Preface

The papers in this volume were presented at the Seventh Annual International Computing and Combinatorics Conference (COCOON 2001), held August 20-23, 2001, in Guilin, China. The topics cover most aspects of theoretical computer science and combinatorics related to computing. The conference was held in cooperation with the Chinese Academy of Sciences and China Computer Federation.

Submissions to the conference this year were conducted entirely online. A total of 97 papers were submitted in time to be considered, of which 50 regular papers and 16 short papers were selected by an international program committee consisting of Eric Allender, Bernard Chazelle, Danny Chen, Jianer Chen, Francis Chin, Kyung-Yong Chwa, Rod Downey, Erich Grädel, Yuri Gurevich, Steven Homer, Toshihide Ibaraki, Tao Jiang, Ker-I Ko, D.T. Lee, Xuemin Lin, Maurice Nivat, R. Ravi, Rüdiger Reischuk, Seinosuke Toda, Jie Wang, Lusheng Wang, Guoliang Xue, and Mihalis Yannakakis.

The authors of submitted papers come from the following countries and regions: Australia, Austria, Bangladesh, Canada, China (including Hong Kong and Taiwan), Czech Republic, France, Germany, India, Israel, Italy, Japan, Korea, New Zealand, The Netherlands, Poland, Russia, Singapore, Spain, Switzerland, U.K., and U.S.A. Each paper was given to at least three Program Committee members, who in some cases were assisted by subreferees. In addition to the selected papers, the conference also included two invited presentations by Bernard Chazelle and Avi Wigderson.

To promote young researchers, the Hao Wang Award this year was given to a paper selected from papers written solely by authors who, at the time of submission, were either students or had received their doctoral degrees within the previous five years. I am happy to announce that the recipient of this award was Xiang-Yang Li for his paper “Generating Well-Shaped d-Dimensional Delaunay Meshes”.

I would like to thank the program committee co-chair Yuri Gurevich for his strong support and all program committee members, their support staff, and subreferees for their excellent work within demanding time constraints. I would also like to thank all authors for submitting their papers to the conference. I am grateful to Steve Tate for letting me use the ACM SIGACT electronic submission service, Minghui Li for helping me create conference web pages, and Richard Cheek for providing system support. Finally, I would like to express my gratitude to Ding-Zhu Du, Xudong Hu, and all local organizers for their hard work in making this meeting possible and enjoyable.

August 2001

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## Table of Contents

### Complexity Theory

Complete Problems for Valiant’s Class of \(qp\)-Computable Families of Polynomials .......................... 1  
*Markus Bläser*

Log-Space Constructible Universal Traversal Sequences for Cycles of Length \(O(n^{4.03})\) ................................. 11  
*Míchal Koucký*

On Universally Polynomial Context-Free Languages .................... 21  
*Nicholas Tran*

Separating Oblivious and Non-oblivious BPs .............................. 28  
*Kazuo Iwama, Yasuo Okabe, Toshiro Takase*

Program Schemes, Queues, the Recursive Spectrum and Zero-One Laws... 39  
*Iain A. Stewart*

Algebraic Properties for P-Selectivity .................................. 49  
*Lane A. Hemaspaandra, Harald Hempel, Arfst Nickelsen*

Parallelizability of Some P-Complete Geometric Problems in the EREW-PRAM .................................................. 59  
*Carla Denise Castanho, Wei Chen, Koichi Wada, Akihiro Fujiwara*

### Computational Biology

Enhanced Sequence Reconstruction with DNA Microarray Application .... 64  
*Samuel A. Heath, Franco P. Preparata*

Non-approximability of Weighted Multiple Sequence Alignment ........... 75  
*Bodo Siebert*

A Greedy Algorithm for Optimal Recombination .......................... 86  
*Shiquan Wu, Xun Gu*

### Computational Geometry

Generating Well-Shaped \(d\)-dimensional Delaunay Meshes .................. 91  
*Xiang-Yang Li*

Towards Compatible Triangulations ...................................... 101  
*Oswin Aichholzer, Franz Aurenhammer, Hannes Krasser, Ferran Hurtado*
<table>
<thead>
<tr>
<th>Title</th>
<th>Page</th>
</tr>
</thead>
<tbody>
<tr>
<td>An Improved Upper Bound on the Size of Planar Convex-Hulls</td>
<td>111</td>
</tr>
<tr>
<td>Abdullah N. Arslan, Ömer Egecioğlu</td>
<td></td>
</tr>
<tr>
<td>On the Planar Two-Watchtower Problem</td>
<td>121</td>
</tr>
<tr>
<td>Sergei Bespamyatnikh, Zhixiang Chen, Kanliang Wang, Binhai Zhu</td>
<td></td>
</tr>
<tr>
<td>Efficient Generation of Triconnected Plane Triangulations</td>
<td>131</td>
</tr>
<tr>
<td>Shin-ichi Nakano</td>
<td></td>
</tr>
<tr>
<td>Packing Two Disks into a Polygonal Environment</td>
<td>142</td>
</tr>
<tr>
<td>Prosenjit Bose, Pat Morin, Antoine Vigneron</td>
<td></td>
</tr>
<tr>
<td>Maximum Red/Blue Interval Matching with Applications</td>
<td>150</td>
</tr>
<tr>
<td>Danny Z. Chen, Xiaobo (Sharon) Hu, Xiaodong Wu</td>
<td></td>
</tr>
<tr>
<td>Computing Farthest Neighbors on a Convex Polytope</td>
<td>159</td>
</tr>
<tr>
<td>Otfried Cheong, Chan-Su Shin, Antoine Vigneron</td>
<td></td>
</tr>
<tr>
<td>Finding an Optimal Bridge between Two Polygons</td>
<td>170</td>
</tr>
<tr>
<td>Xuehou Tan</td>
<td></td>
</tr>
<tr>
<td>How Good Is Sink Insertion?</td>
<td>181</td>
</tr>
<tr>
<td>Xiang-Yang Li, Yu Wang</td>
<td></td>
</tr>
<tr>
<td>Polynomial Time Algorithms for Three-Label Point Labeling</td>
<td>191</td>
</tr>
<tr>
<td>Rob Duncan, Jianbo Qian, Binhai Zhu</td>
<td></td>
</tr>
<tr>
<td>Approximation Algorithms for the Watchman Route and Zookeeper’s Problems</td>
<td>201</td>
</tr>
<tr>
<td>Xuehou Tan</td>
<td></td>
</tr>
</tbody>
</table>

**Data Structures and Algorithms**

<table>
<thead>
<tr>
<th>Title</th>
<th>Page</th>
</tr>
</thead>
<tbody>
<tr>
<td>PC-Trees vs. PQ-Trees</td>
<td>207</td>
</tr>
<tr>
<td>Wen-Lian Hsu</td>
<td></td>
</tr>
<tr>
<td>Stacks versus Deques</td>
<td>218</td>
</tr>
<tr>
<td>Holger Petersen</td>
<td></td>
</tr>
<tr>
<td>Optimizing a Computational Method for Length Lower Bounds for Reflecting Sequences</td>
<td>228</td>
</tr>
<tr>
<td>H.K. Dai</td>
<td></td>
</tr>
</tbody>
</table>

**Games and Combinatorics**

<table>
<thead>
<tr>
<th>Title</th>
<th>Page</th>
</tr>
</thead>
<tbody>
<tr>
<td>Competitive Facility Location along a Highway</td>
<td>237</td>
</tr>
<tr>
<td>Hee-Kap Ahn, Siu-Wing Cheng, Otfried Cheong, Mordecai Golin, René van Oostrum</td>
<td></td>
</tr>
</tbody>
</table>
Membership for Core of LP Games and Other Games...................... 247
Qizhi Fang, Shanfeng Zhu, Maocheng Cai, Xiaotie Deng

Strong Solutions to the Identification Problem ......................... 257
Pino Caballero-Gil, Candelaria Hernández-Goya

Area Efficient Exponentiation
Using Modular Multiplier/Squarer in GF($2^m$) ......................... 262
Hyun-Sung Kim, Kee-Young Yoo

Graph Algorithms and Complexity

A Space Saving Trick for Directed Dynamic Transitive Closure
and Shortest Path Algorithms ......................................... 268
Valerie King, Mikkel Thorup

Finding the Most Vital Node of a Shortest Path ....................... 278
Enrico Nardelli, Guido Proietti, Peter Widmayer

Algorithm for the Cost Edge-Coloring of Trees ....................... 288
Xiao Zhou, Takao Nishizeki

Counting $H$-Colorings of Partial $k$-Trees .......................... 298
Josep Díaz, Maria Serna, Dimitrios M. Thilikos

A Linear Time Algorithm for Enumerating All the Minimum and Minimal
Separators of a Chordal Graph ...................................... 308
L. Sunil Chandran

Graph Separators: A Parameterized View .............................. 318
Jochen Alber, Henning Fernau, Rolf Niedermeier

On Assigning Prefix Free Codes to the Vertices of a Graph .......... 328
N.S. Narayanaswamy, C.E. Veni Madhavan

A New Measure of Edit Distance Between Labeled Trees ............. 338
Chin Lung Lu, Zheng-Yao Su, Chuan Yi Tang

A Highly Efficient Algorithm to Determine Bicritical Graphs ......... 349
Dingjun Lou, Ning Zhong

Graph Drawing

Layered Drawings of Graphs with Crossing Constraints ............... 357
Irene Finocchi

On the Validity of Hierarchical Decompositions ........................ 368
Irene Finocchi, Rossella Petreschi
Graph Theory

Lower Bounds on the Minus Domination and \( k \)-Subdomination Numbers .. 375
Liying Kang, Hong Qiao, Erfang Shan, Ding-Zhu Du

Edge Connectivity vs Vertex Connectivity in Chordal Graphs ............ 384
L. Sunil Chandran

Changing the Diameter of Graph Products .............................. 390
Ting-Yi Sung, Jeng-Jung Wang

Plane Graphs with Acyclic Complex .................................... 395
Baogang Xu

On the Domination Numbers of Generalized de Bruijn Digraphs and Generalized Kautz Digraphs ................................. 400
Yosuke Kikuchi, Yukio Shibata

A Notion of Cross-Perfect Bipartite Graphs ............................... 409
Milind Dawande

Some Results on Orthogonal Factorizations ................................ 414
Haodi Feng

Cluttered Orderings for the Complete Graph ................................ 420
Myra B. Cohen, Charles J. Colbourn, Dalibor Froncek

Online Algorithms

Improved On-Line Stream Merging:
From a Restricted to a General Setting ................................. 432
Wun-Tat Chan, Tak-Wah Lam, Hing-Fung Ting, Wai-Ha Wong

On-Line Deadline Scheduling on Multiple Resources ..................... 443
Jae-Hoon Kim, Kyung-Yong Chwa

Competitive Online Scheduling with Level of Service .................... 453
Ée-Chien Chang, Chee Yap

On-Line Variable Sized Covering ......................................... 463
Leah Epstein

Randomized and Average-Case Algorithms

On Testing for Zero Polynomials by a Set of Points with Bounded Precision ................................. 473
Jin-Yi Cai, Eric Bach

A Randomized Algorithm for Gossiping in Radio Networks ............ 483
Marek Chrobak, Leszek Gąsieniec, Wojciech Rytter
Deterministic Application of Grover’s Quantum Search Algorithm........ 493
Kyoichi Okamoto, Osamu Watanabe

Random Instance Generation for MAX 3SAT............................. 502
Mitsuo Motoki

**Steiner Trees**

The Euclidean Bottleneck Steiner Tree and Steiner Tree with Minimum Number of Steiner Points......................... 509
Dingzhu Du, Lusheng Wang, Baogang Xu

An FPTAS for Weight-Constrained Steiner Trees in Series-Parallel Graphs.................................................. 519
Guangting Chen, Guoliang Xue

**Systems Algorithms and Modeling**

Decidable Approximations on Generalized and Parameterized Discrete Timed Automata........................................ 529
Zhe Dang, Oscar H. Ibarra, Richard A. Kemmerer

Multiplicative Adaptive Algorithms for User Preference Retrieval ........ 540
Zhixiang Chen

Parametric Scheduling for Network Constraints............................... 550
K. Subramani

A Logical Framework for Knowledge Sharing in Multi-agent Systems ..... 561
Kaile Su, Xudong Luo, Huaiqing Wang, Chengqi Zhang, Shichao Zhang, Qingfeng Chen

A Lockout Avoidance Algorithm without Using Time-Stamps for the $k$-Exclusion Problem....................................... 571
Kumiko Obokata, Michiko Omori, Kazuhiro Motegi, Yoshihide Igarashi

**Computability**

Prefix-Free Languages and Initial Segments
of Computably Enumerable Degrees........................................ 576
Guohua Wu

Weakly Computable Real Numbers
and Total Computable Real Functions...................................... 586
Robert Rettinger, Xizhong Zheng, Romain Gengler,
Burchard von Braunmühl

Turing Computability of a Nonlinear Schrödinger Propagator........ 596
Klaus Weihrauch, Ning Zhong

**Author Index**.......................................................... 601
Complete Problems for Valiant’s Class of qp-Computable Families of Polynomials

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Abstract. We prove that the families matrix powering, iterated matrix product, and adjoint matrix are VQP–complete, where VQP denotes Valiant’s class of quasipolynomial–computable families of multivariate polynomials. This proves a conjecture by Bürgisser [3, Conjecture 8.1].

1 Introduction

In connection with his famous #P–completeness result [11] for the permanent, Valiant presented an algebraic analogue of the theory of NP–completeness [10] (a decade before the proposal of Blum, Shub, and Smale [2]). In the present work, we prove the completeness of various families of multivariate polynomials for Valiant’s algebraic class of qp–computable families of multivariate polynomials. For the reader’s convenience, we first present the relevant details of Valiant’s theory. More material can be found in [4,3].

A function $t : \mathbb{N} \rightarrow \mathbb{N}$ is called $p$–bounded if $t(n) \leq O(n^c)$ for some constant $c$. The function $t$ is called $p$–bounded from above and below if in addition $t(n) \geq \Omega(n^{1/c})$. A function $u : \mathbb{N} \rightarrow \mathbb{N}$ is called qp–bounded if $u(n) \leq n^{O(\log^c n)}$.

The main objects in Valiant’s algebraic NP–completeness theory are certain families of multivariate polynomials. Throughout the remainder of this work, $k$ denotes a field.

Definition 1. A sequence $f = (f_n)$ of multivariate polynomials over $k$ is called a $p$–family if the number of variables and the degree of $f_n$ are both $p$–bounded functions in $n$.

The complexity $L(f)$ of a multivariate polynomial $f \in k[X_1, \ldots, X_m]$ is the size of a smallest arithmetic circuit over the basis $\{+,-,\ast\}$ (all operations with fanin two) that computes $f$ from the indeterminates $X_1, \ldots, X_m$ and the scalars in $k$.

Definition 2. 1. A $p$–family $f = (f_n)$ is $p$–computable if the complexity $L(f_n)$ is a $p$–bounded function in $n$. The class of all $p$–computable families is denoted by $\text{VP}$ (or $\text{VP}_k$).

2. The $p$–family $f$ is qp–computable if $L(f_n)$ is a qp–bounded function in $n$.

The class of all qp–computable families is denoted by $\text{VQP}$ (or $\text{VQP}_k$).
In the same way, we can also define the class \( VNP \), an algebraic analogue of \( NP \). Since a detailed definition of \( VNP \) is more complicated than those of the classes \( VP \) and \( VQP \), the reader is referred to [4,3]. The following hypotheses play a central role in Valiant’s theory—like Cook’s Hypothesis in structural complexity theory.

**Hypothesis 1.**

1. Valiant’s Hypothesis: \( VP \subsetneq VNP \) over any field.
2. Extended Valiant’s Hypothesis: \( VNP \setminus VQP \neq \emptyset \) over any field.

In Valiant’s model, families of polynomials are compared by a simple type of reductions, namely projections. A polynomial \( f(X_1, \ldots, X_m) \) is called a projection of another polynomial \( g(Y_1, \ldots, Y_n) \) if there are \( s_1, \ldots, s_n \in k \cup \{ X_1, \ldots, X_m \} \) such that

\[
    f(X_1, \ldots, X_m) = g(s_1, \ldots, s_n).
\]

In other words, \( f \) is obtained from \( g \) by replacing each of the \( Y_1, \ldots, Y_n \) by a scalar from \( k \) or by one of the \( X_1, \ldots, X_m \).

**Definition 3.**

1. A family \( f = (f_n) \) is a \( p \)-projection of \( g = (g_m) \), in symbols \( f \leq_p g \), if there exists a function \( r \) that is \( p \)-bounded from above and below such that

\[
    \exists n_0 \forall n \geq n_0 : f_n \text{ is a projection of } g_{r(n)}.
\]

2. A family \( f = (f_n) \) is a \( qp \)-projection of \( g = (g_m) \), in symbols \( f \leq_{qp} g \), if there exists a \( qp \)-bounded function \( r \) that is also \( p \)-bounded from below such that

\[
    \exists n_0 \forall n \geq n_0 : f_n \text{ is a projection of } g_{r(n)}.
\]

The relations “is a \( p \)-projection of” and “is a \( qp \)-projection of” are both transitive. Obviously, if \( g \leq_p f \), then also \( g \leq_{qp} f \).

The next step is to define \( VQP \)-complete families.

**Definition 4.** A \( qp \)-computable family \( f \) is called \( VQP \)-complete if \( g \leq_{qp} f \) for all \( g \in VQP \).

As expected, if \( f \) is \( VQP \)-complete and \( f \leq_{qp} h \), then \( h \) is also \( VQP \)-complete.

Up to now, the only known \( VQP \)-complete family is the determinant family defined by

\[
    DET_n = \det(X_{i,j})_{1 \leq i,j \leq n}.
\]

(For a proof, see [3, Corollary 2.29].) As our main results, we exhibit further \( VQP \)-complete families of multivariate polynomials. Specifically, we show that the families iterated matrix multiplication \( IMM \), matrix powering \( POW \), and adjoint matrix \( ADJ \) defined by

\[
    IMM_n = \text{trace}(X_1 \cdots X_n),
\]
\[
    POW_n = \text{trace}(X^{[n/3]}),
\]
\[
    ADJ_n = \text{trace}((\det(X) \cdot X^{-1})
\]

are all \( VQP \)-complete. (Above, \( X \) and the \( X_{\nu} \) denote \( n \times n \)-matrices with independent indeterminate entries.) This proves a conjecture by Bürgisser [3, Conjecture 8.1]. The original conjecture for \( POW \) made by Bürgisser is that \( (\det(X^n)) \) is \( VQP \)-complete. This remains an open problem.
The main technical problems we are confronted with are that we have a very weak type of reductions, namely $qp$–projections, and that we only compute one result.

For any constant $c \in \mathbb{N}$, let the family \textit{constant size $c$ iterated matrix multiplication $c$–CIMM} be defined through

$$c$–CIMM$ _n = \text{trace}(Y_1 \cdots Y_n),$$

where the $Y_\nu$ are $c \times c$–matrices (and not $n \times n$–matrices) with independent indeterminate entries. Strengthening the above $VQP$–completeness result for IMM, we also show that for any $c \geq 3$, the family $c$–CIMM is $VQP$–complete. This follows from the results by Ben-Or and Cleve [1].

2 Auxiliary Results

We briefly review some (well known) results, which we will need later on, and provide some useful variations of these auxiliaries. For a multivariate polynomial $f \in k[X_1, \ldots, X_m]$ the depth $D(f)$ is the minimal depth of an arithmetic circuit over the basis $\{+, -, *\}$ (all operations with fanin two) that computes $f$ from the indeterminates $X_1, \ldots, X_m$ and the scalars in $k$.

\textbf{Theorem 1 (Hyafil, Valiant et al.).} \textit{If $f$ is an $n$–variate polynomial of degree $d \geq 1$, then}

$$D(f) \leq O(\log(d \cdot L(f)) \log d + \log n).$$

\textit{Moreover, there is an arithmetic circuit of depth $O(\log(d \cdot L(f)) \log d + \log n)$ and size $d^6L(f)^3$ that computes $f$.}

For a proof of the first part of the statement, see [7]. It can also be shown that the constant hidden in the $O$–notation is universal. The second claim is due to Valiant et al. [9]. The reader is also referred to [4, Thm. 21.35]. Miller et al. [8] present an online construction.

For technical reasons, we introduce the family $c$–CIMM$'$ defined through

$$c$–CIMM$'_n = \text{the entry in position (1, 1) of } Y_1 \cdots Y_n,$$

where the $Y_\nu$ are $c \times c$–matrices with independent indeterminate entries. The family $3$–CIMM$'$ plays an important role for our completeness results.

\textbf{Theorem 2 (Ben-Or & Cleve).} \textit{Let $f$ be a multivariate polynomial. Moreover, let $\ell = 4D(f)$. Then $f$ is a projection of $3$–CIMM$'_\ell$.}

For a proof, see [1, Thm. 1]. Since Ben-Or and Cleve focused on computing polynomial size formulas with polynomial length straight-line programs that use a constant number of registers, they only stated the above theorem for formulas. Nevertheless, Ben-Or and Cleve’s approach also works for arbitrary straight-line programs, since a straight-line program of depth $d$ induces a formula of depth $d$ and size at most $2^d$. 
Since

\[(1, 1)\text{-entry of } Y_1 \cdots Y_n = \text{trace}(E_{1,1} \cdot Y_1 \cdots Y_n \cdot E_{1,1}),\]

where \(E_{1,1}\) denotes the \(c \times c\)–matrix that has a one in position \((1, 1)\) and zeros elsewhere, \(c\text{–CIMM}_n'\) is a projection of \(c\text{–CIMM}_{n+2}\). Therefore, the below theorem follows.

**Theorem 3.** Let \(f\) be a multivariate polynomial and let \(\ell = 4^{D(f)} + 2\). Then \(f\) is a projection of \(3\text{–CIMM}_\ell\).

### 3 Completeness Results

In this section, we prove as our main result the \(VQP\text{–completeness of the families } IMM, POW, \text{ and } ADJ\). Since each of these three families is in \(VP\), it suffices to show the hardness.

#### 3.1 Iterated Matrix Multiplication

**Theorem 4.** For any \(c \geq 3\), \(c\text{–CIMM}\) is \(VQP\text{–complete (under qp–projections). The same is true for } c\text{–CIMM}_n'\).

**Proof.** The polynomial \(c\text{–CIMM}_n\) can be computed by an arithmetic circuit of size \(O(n)\). Thus \(c\text{–CIMM} \in VP \subseteq VQP\).

Since obviously \(c\text{–CIMM}_n\) is a projection of \(c'\text{–CIMM}_n\) for each \(n\) and each \(c' \geq c\), it is sufficient to show that for any \(f = (f_n) \in VQP\), \(f \leq_{qp} 3\text{–CIMM}\). By definition, we know that the number of variables and the degree of \(f_n\) are p–bounded functions and that \(L(f_n)\) is a qp–bounded function. By Theorem 1, \(D(f_n) \leq O(\log^m n)\) for some constant \(m\). By Theorem 3, \(f_n\) is a projection of \(3\text{–CIMM}_\ell\) where \(\ell = n^{O(\log^m n)}\). Hence \(f\) is a qp–projection of \(3\text{–CIMM}\).

The result for \(c\text{–CIMM}_n'\) follows completely alike.

Since \(3\text{–CIMM}_n\) is a projection of \(IMM_n\), we immediately obtain the following corollary.

**Corollary 1.** The family \(IMM\) is \(VQP\text{–complete (under qp–projections).}\)

The next corollary is of rather technical nature. It describes a “self reducibility” property of \(c\text{–CIMM}\). We will rely on the corollary in the subsequent sections.

**Corollary 2.** Let \(c \geq 3\) and let \(i\) and \(j\) be functions \(\mathbb{N} \to \{1, 2, \ldots, c\}\). Furthermore, let \((\alpha_n)\) and \((\beta_n)\) be sequences of scalars from \(k\) and define \(w(n) = \min\{i \geq n \mid \alpha_i \neq 0\}\). If \(w\) is qp–bounded, then \((\alpha_n \cdot f_n + \beta_n)\) is \(VQP\text{–complete, where } f_n\) is the entry in position \((i(n), j(n))\) of a generic iterated \(3 \times 3\text{–matrix product of length } n\).
Proof. We first prove that \((c - CIMM'_n + \beta_n)\) is VQP–complete. Then we reduce \((c - CIMM'_n + \beta_n)\) to \((\alpha_n \cdot f_n + \beta_n)\).

For the first part, let \((g_n) \in \text{VQP}\) be given. Then also \((g_n - \beta_n) \in \text{VQP}\). By the VQP–completeness of \(c - CIMM'\), \((g_n - \beta_n)\) is a qp–projection of \(c - CIMM'\). Thus \((g_n)\) is a qp–projection of \((c - CIMM'_n + \beta_n)\). Since \((g_n)\) was arbitrary, this shows the VQP–completeness of \((c - CIMM'_n + \beta_n)\).

Now we reduce \((c - CIMM'_n + \beta_n)\) to \((\alpha_n \cdot f_n + \beta_n)\). To obtain \(c - CIMM'_n + \beta_n\), we substitute \(w(n+3) - n - 3\) many copies of the \(c \times c\)–identity matrix and one \(\alpha_{w(n+3)}^{-1}\) multiple of the \(c \times c\)–identity matrix into \(\alpha_{w(n+3)} \cdot f_{w(n+3)} + \beta_{w(n+3)}\) at arbitrary places except the first and the last matrix. (Note that \(w(\nu) \geq \nu\) for all \(\nu\).) As the first and the last matrix, we substitute appropriate permutation matrices that move the entry in position \((i(n), j(n))\) into position \((1, 1)\). Since \(w\) is qp–bounded, \((c - CIMM'_n + \beta_n)\) is a qp–projection of \((\alpha_n \cdot f_n + \beta_n)\). \(\Box\)

3.2 Matrix Powering

Our plan is to modify the reduction to iterated matrix multiplication given by Cook \([5]\). The weak type of reductions however is a major obstacle.

Let \(X_1, \ldots, X_m\) be \(3 \times 3\)–matrices with independent indeterminate entries. Let \(Y\) be the block matrix of size \(3m \times 3m\)

\[
Y = \begin{pmatrix}
0 & X_1 & 0 & \cdots & 0 \\
0 & 0 & X_2 & 0 \\
\vdots & \ddots & \ddots & \ddots \\
0 & 0 & 0 & \cdots & X_{m-1} \\
X_m & 0 & 0 & \cdots & 0
\end{pmatrix}.
\]

(Note the extra \(X_m\) in the lower left corner. Without it, all powers of \(Y\) would have trace zero.) A straightforward calculation shows that

\[
Y^m = \begin{pmatrix}
P_1 & 0 & 0 & \cdots & 0 \\
0 & P_2 & 0 & 0 \\
\vdots & \ddots & \ddots & \ddots \\
0 & 0 & 0 & \cdots & 0 \\
0 & 0 & 0 & \cdots & P_m
\end{pmatrix},
\]

(1)

where \(P_\mu = X_\mu X_{\mu+1} \cdots X_m X_1 \cdots X_{\mu-1}\). By the below fact,

\[
\text{trace}(Y^m) = m \cdot \text{trace}(X_1 \cdots X_m) = m \cdot 3 \cdot \text{CIMM}_m.
\]

Fact 1. 1. For any two \(\ell \times \ell\)–matrices \(A\) and \(B\), \(\text{trace}(AB) = \text{trace}(BA)\).

2. If we write \(A = (A_{i,j})_{1 \leq i,j \leq q}\) as a block matrix with \((\ell/q) \times (\ell/q)\)–blocks \(A_{i,j}\), then \(\text{trace} A = \text{trace} A_{1,1} + \text{trace} A_{2,2} + \cdots + \text{trace} A_{q,q}\).
Thus \( m \cdot 3 - C I M M_m \) is a projection of \( P O W_{3m} \). If the characteristic of \( k \) is zero, then \( m \cdot 3 - C I M M_m \) is always nonzero. If the characteristic \( p > 0 \), then at least every second of the \( m \cdot 3 - C I M M_m \) is nonzero. Thus the function \( j \mapsto \min\{i \geq j \mid i \not\equiv 0 \ (p)\} \) is \( q p \)-bounded. Hence by Corollary 2, the family \( (m \cdot 3 - C I M M_m) \) is \( V Q P \)-complete and so is \( P O W \).

We define the family \( m - P O W \) for a given function \( m : \mathbb{N} \to \mathbb{N} \) by

\[
m - P O W_n = \text{trace}(X^{m(n)}),
\]

where \( X \) denotes a matrix with independent indeterminate entries of size \( n \times n \).

For functions \( m \) such that \( m(n) \leq \lfloor n/3 \rfloor \) for all \( n \) and \( m \) is \( p \)-bounded from below, the family \( m - P O W \) is \( V Q P \)-complete (under \( q p \)-projections).

### 3.3 Matrix Inversion

Let \( X_1, \ldots, X_n \) be \( 3 \times 3 \)–matrices with indeterminate and independent entries. Let \( N \) be the \( 3(n+1) \times 3(n+1) \)–matrix

\[
N = \begin{pmatrix}
0 & X_1 & 0 & \cdots & 0 \\
0 & 0 & X_2 & 0 & \cdots \\
0 & 0 & 0 & \cdots & X_n \\
0 & 0 & 0 & \cdots & 0
\end{pmatrix}
\]  

The matrix \( N \) is nilpotent and we have

\[
(I - N)^{-1} = I + N + N^2 + \cdots + N^n = \begin{pmatrix}
I & P_{1,1} & P_{1,2} & \cdots & P_{1,n} \\
0 & I & P_{2,2} & P_{2,n} & \cdots \\
0 & 0 & I & \cdots & P_{n,n} \\
0 & 0 & 0 & \cdots & I
\end{pmatrix},
\]

where \( I \) denotes the identity matrix of appropriate size (\( 3n + 3 \) on the left-hand and \( 3 \) on the right-hand side) and \( P_{i,j} \) stands for the product \( X_i \cdots X_j \) for \( i \leq j \).
Let $S$ be the permutation matrix

$$
S = \begin{pmatrix}
0 & I & 0 & \cdots & 0 \\
0 & 0 & I & 0 & \cdots \\
\vdots & \ddots & \ddots & \ddots & \ddots \\
0 & 0 & 0 & \cdots & I \\
I & 0 & 0 & \cdots & 0
\end{pmatrix}.
$$

We have

$$(S^{-1}(I - N))^{-1} = (I - N)^{-1}S = \begin{pmatrix}
P_{1,n} & P_{1,1} & \cdots & P_{1,n-1} \\
P_{2,n} & 0 & I & P_{2,n-2} \\
\vdots & \ddots & \ddots & \ddots \\
P_{n,n} & 0 & 0 & \cdots & I \\
I & 0 & 0 & \cdots & 0
\end{pmatrix}.
$$

(3)

Thus

$$
\text{trace}(S^{-1}(I - N))^{-1} = \text{trace} P_{1,n} = 3 - CIMM_n.
$$

Because $\det(S^{-1}(I - N)) = \pm 1$, $\pm 3 - CIMM_n$ is a projection of $ADJ_{3(n+1)}$.

(4) This yields the following completeness result.

**Theorem 6.** The family $ADJ$ is VQP-complete (under qp-projections).

4 Further Directions

What happens if we do not take the trace in the definition of the families $IMM$, $POW$, and $ADJ$, but look at the following families given by

$$
IMM'_n = (1,1)\text{-entry of } X_1 \cdots X_n,
$$

$$
POW'_n = (1,1)\text{-entry of } X_{\lfloor n/3 \rfloor},
$$

$$
ADJ'_n = (1,1)\text{-entry of } \det(X) \cdot X^{-1}?
$$

From Theorem 4, (1), and (3), respectively, it follows that these families are also VQP-complete. Thus we have the following theorem.

**Theorem 7.** The above families $IMM'$, $POW'$, and $ADJ'$ are VQP-complete (under qp-projections).

**Remark 1.** By applying appropriate permutation matrices, we could also define $IMM'$ and $ADJ'$ not by taking the $(1,1)$-entry but any other entry and would still get the same completeness results for these families. Since Valiant’s model is nonuniform, the entries may also vary throughout the whole sequence.
Can we achieve a similar result for the family $POW^f$? In fact, we can. The proof is however a little more elaborate.

**Theorem 8.** For any functions $i,j : \mathbb{N} \to \mathbb{N}$ with $1 \leq i(n), j(n) \leq n$ for all $n$, the family $((i(n), j(n))$–entry of $X^{\lfloor n/3 \rfloor})$ is a VQP–complete family (under qp–projections).

**Proof.** Let w.l.o.g. $n = 3m$. We distinguish the cases $i(n) = j(n)$ and $i(n) \neq j(n)$. In the first case, note that in (1), each entry on the diagonal corresponds to an entry of an iterated product of $3 \times 3$–matrices of length $m$.

For the case $i(n) \neq j(n)$, note that also the entry in position $(1,2)$ of $Y^n$ in (1) is an entry of an iterated $3 \times 3$–matrix product. Let $P$ be the permutation matrix that exchanges the first with the $i(n)$th and the second with the $j(n)$th row resp. column. The matrix $P$ is its own inverse. We now consider $PYP$ instead of $Y$ in (1). We have $(PYP)^m = PY^m P$. By the choice of $P$, the entry in position $(i,j)$ of $PY^m P$ is an entry of an iterated $3 \times 3$–matrix product.

Altogether, we obtain a VQP–complete family. $\square$

Over fields of characteristic zero, we can even prove the VQP–completeness of the family $((1,1)$–entry of $X^n)$ corresponding to the $n$–powering problem, a task which remains unsolved in the case of taking the trace.

**Theorem 9.** Let $k$ be a field of characteristic zero. For any $i,j : \mathbb{N} \to \mathbb{N}$ with $1 \leq i(n), j(n) \leq n$ for all $n$ and any polynomial $p$, $((i(n), j(n))$–entry of $X^{p(n)})$ is a VQP–complete family (under qp–projections).

**Proof.** Since $p$ is a polynomial, $f$ is in $VP \subseteq VQP$.

It remains to show the hardness: let $X_1, \ldots, X_n$ be $3 \times 3$–matrices and define $N$ like in (2). The matrix $N$ is nilpotent, specifically, $N^{n+1} = 0$. We have

$$(I + N)^{p(n)} = \sum_{\nu=0}^{n} \binom{p(n)}{\nu} N^\nu$$

$$= \begin{pmatrix} I \cdot p(n) \cdot P_{1,1} & \binom{p(n)}{2} \cdot P_{1,2} & \cdots & \binom{p(n)}{n} \cdot P_{1,n} \\ 0 & I \cdot p(n) \cdot P_{2,2} & \cdots & \binom{p(n)}{n-1} \cdot P_{2,n} \\ \vdots & \ddots & \ddots & \vdots \\ 0 & 0 & \cdots & p(n) \cdot P_{n,n} \\ 0 & 0 & \cdots & I \end{pmatrix},$$

where $I$ denotes the identity matrix of appropriate size and $P_{i,j}$ stands for the product $X_i \cdots X_j$.

If $i(n) \neq j(n)$, then we can move the entry in position $(1,n)$ to position $(i(n), j(n))$ by switching over from $(I + N)$ to $P(I + N)P$ for an appropriate permutation matrix (like in the proof of Theorem 3). After that, the entry in position $(i(n), j(n))$ is the $\binom{p(n)}{n}$–multiple of an entry of an iterated $3 \times 3$–matrix product.
It remains the case $i(n) = j(n)$. We may assume that $i(n) = j(n) = 1$. Otherwise, we switch over to $P(I + N)P$ where $P$ is the permutation matrix that exchanges the first and the $i(n)$th row resp. column. Let $T$ be the matrix

$$T = \begin{pmatrix}
I & 0 & 0 & \cdots & 0 \\
0 & I & 0 & & 0 \\
\vdots & \ddots & \ddots & \ddots & \vdots \\
0 & 0 & 0 & \ddots & 0 \\
I & 0 & 0 & \cdots & I
\end{pmatrix}.$$ 

The inverse of $T$ coincides with $T$ except that the identity matrix in the lower left corner has to be replaced by $-I$. A straight forward calculation shows that the $3 \times 3$–matrix in the upper left block of $(T^{-1}YT)^{p(n)} = T^{-1}Y^{p(n)}T$ equals $(p(n))_n \cdot P_{1,n} + I$. Thus the entry in position $(1, 1)$ of $(T^{-1}YT)^{p(n)}$ is just $(p(n))_n \cdot 3 \cdot \text{CIMM}'_n + 1$.

Since the characteristic of $k$ is zero, the coefficient $(p(n))_n \neq 0$. Together with Corollary 2 this proves the claim of the theorem.

Once we have proven the completeness of $\text{IMM}$, $\text{POW}$, and $\text{ADJ}$, many other families of polynomials corresponding to problems from linear algebra turn out to be $\text{VQP}$–complete.

As an example, we examine the family $\text{LINEQ}$

$$\text{LINEQ}_n = \text{first entry of } \det(X) \cdot X^{-1}b,$$

which corresponds to solving a system of linear equations $Xa = b$. Above, $X$ denotes an $n \times n$–matrix and $b$ a (column) vector of length $n$ with independent indeterminate entries.

If we specialize $b_1 \mapsto 1$ and $b_\nu \mapsto 0$ for $2 \leq \nu \leq n$, then the image of $\text{LINEQ}_n$ under this specialization is $\text{ADJ}'_n$. Thus, the family $\text{LINEQ}$ is $\text{VQP}$–complete.

**Theorem 10.** The family $\text{LINEQ}$ is $\text{VQP}$–complete (under qp–projections).

Since the permanent family $\text{PER}$ is $\text{VNP}$–complete over fields of characteristic other than two, the Extended Valiant’s Hypothesis is equivalent to the following completely algebraic statements by our $\text{VQP}$–completeness results: the Extended Valiant’s Hypothesis holds iff $\text{PER}$ is not a qp–projection of $\text{IMM}$ nor of $\text{POW}$ nor of $\text{ADJ}$ nor of any other of the $\text{VQP}$–complete families considered in this work.

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References

Log-Space Constructible Universal Traversal Sequences for Cycles of Length $O(n^{4.03})$
(Extended Abstract)

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Abstract. The paper presents a simple construction of polynomial length universal traversal sequences for cycles. These universal traversal sequences are log-space (even $NC^1$) constructible and are of length $O(n^{4.03})$. Our result improves the previously known upper-bound $O(n^{4.76})$ for log-space constructible universal traversal sequences for cycles.

1 Introduction

One of the major problems in computer science is the graph $s,t$-connectivity problem. It is well known that a directed version of this problem is complete for nondeterministic log-space ($NL$), and its undirected version is complete for symmetric log-space ($SL$). It has been conjectured that symmetric log-space is equal to deterministic log-space ($SL = L$).

One approach to proving that conjecture is to use universal traversal sequences introduced by Cook (see [AKL+]). A traversal sequence for a $d$-regular graph is a sequence of numbers from $\{1, \ldots, d\}$, which directs us in traversing the graph. For given $d$ and $n$, a universal traversal sequence for $d$-regular graphs of size $n$ is a traversal sequence which completely traverses any $d$-regular graph of size $n$ starting at any vertex. [AKL+] gave a probabilistic argument for existence of universal traversal sequences of polynomial length in $d$ and $n$. In a sequence of papers [N], [NSW], [ATWZ] it was shown using derandomization of that argument that $SL \subseteq DSPACE(\log^{4/3} n)$.

A more direct approach to proving $SL = L$ is to construct an explicit log-space constructible universal traversal sequence for $d$-regular graphs, where $d \geq 3$. Indeed, explicit constructions of universal traversal sequences for specific classes of graphs exist. For 2-regular graphs (cycles), [BBK+], and [B] constructed universal traversal sequences of length $n^{O(\log n)}$, which were later superseded by log-space constructible universal traversal sequences for cycles of length $O(n^{4.76})$ by Istrail [I]. In this paper we present a simple log-space construction of universal traversal sequences for cycles of length $O(n^{4.03})$ (which we believe can be improved up to $O(n^{3.99})$ by setting parameters of the construction optimally.) The best known lower bound $\Omega(n^{1.43})$ and non-constructive upper

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bound $O(n^3)$ on the length of universal traversal sequences for cycles are due to [BT] and [A], respectively. (See [BT] for comprehensive overview. Recently, it was shown in [K2] that a different kind of traversal sequences does not obey these lower bounds.) Explicit universal traversal sequences of length $n^{O(\log n)}$ for complete graph were presented in [KPS], and [HW] presented construction of length $n^{O(d \log d)}$ for $d$-regular expanders. The best construction of universal traversal sequences for 3-regular graphs is based on Nisan’s pseudo-random generator [N]; these sequences are of length $n^{O(\log n)}$.

1.1 Overview of the Construction

Our construction has two parts:

1. The first part (Section 3) is a construction of an $n^{O(\log n)}$ universal traversal sequence (UTS) for cycles of length $n$. This part uses essentially the same idea as a construction of Bridgland [B]. It is a recursive construction in which at every iteration a UTS for cycles of length $c_n$, $c \geq 2$, is constructed from a UTS for cycles of length $n$. Every iteration consists of two dual stages. The depth of the recursion is $O(\log n)$ and the factor by which the generated universal traversal sequence is expanded at each stage is $O(n)$.

2. The second part (Section 4) involves again two dual stages that reduce the expansion factor to $O(1)$, and they interleave the stages of the first part.

At both stages of the first part the traversed graph is reduced to a smaller one, whereas at both stages of the second part the graph is expanded a little bit by inserting new vertices. After applying all four stages the graph is smaller by at least a constant factor.

The main contribution of this paper lies in the construction of the second part where we show that we can modify the traversed graph by inserting new vertices so that this modification is transparent for traversal sequences.

The previous construction by Istrail was also based on Bridgland’s construction. Istrail’s construction also reduced the expansion factor to a constant as our does but using slightly different (ad hoc) approach. Our construction seems more natural and better motivated hence, much simpler to understand.

2 Preliminaries

Let $G = (V, E)$ be an undirected cycle. Let $V = \{v_1, \ldots, v_n\}$ so that $E = \{(v_i, v_{i+1}); 1 \leq i < n\} \cup \{(v_1, v_n)\}$. We may look at every edge $(u, v)$ in $G$ as a pair of two directed edges $(u, v)$ and $(v, u)$. For $1 \leq i < n$, the directed edges $(v_i, v_{i+1})$ and $(v_n, v_1)$ are called the right edges, and $(v_{i+1}, v_i)$ and $(v_1, v_n)$ are called the left edges. Similarly, $v_{i+1}$ is called a right neighbor of $v_i$, and $v_i$ is a left neighbor of $v_{i+1}$. At every vertex in $G$, let one of the outgoing edges be labeled by 1 and the other one by 0. We label every vertex in $G$ according to the label of its right outgoing edge.

A 0-1 sequence $b(v_1), b(v_2), \ldots, b(v_n)$ uniquely describes $G$, where $b(v_i)$ is the label of vertex $v_i$. We call such a sequence a graph sequence corresponding to $G$. 
Two graph sequences are considered to be the same if one of them is a cyclic shift of the other one. Hence, we consider any graph sequence to be wrapped around (cyclic). We consider all the graphs with the same graph sequence to be the same. We identify every graph sequence with its corresponding graph (actually a class of the same graphs.) We refer to the digits of a graph sequence as vertices. (We use term vertex also for digits of any 0-1 sequence.)

A traversal sequence is a 0-1 sequence. “To run a traversal sequence \( t \) on a graph \( G \) starting at a vertex \( v \)” means to walk in \( G \) starting at \( v \) and following edges labeled consistently with \( t \). Note, if the label of a vertex in the traversal sequence matches the label of the traversed vertex then we move to the right neighbor, otherwise to the left one. Because we identify a graph with its graph sequence we may also talk about running a traversal sequence on a 0-1 sequence.

Let \( s \) be a 0-1 sequence. A 0-run (1-run) in \( s \) is any maximal part of \( s \) consisting only of 0’s (1’s). The type of the run, 0-run or 1-run, is called a color of that run. The leftmost and rightmost vertices of the run are called border vertices. The border vertices of a run of length one coincide. A 0-1 sequence consisting only of runs of length \( i \) and \( j \) is \( (i, j) \)-sequence e.g., 000 111 000 111111 000 is a \((3, 6)\)-sequence. A sequence consisting of a single run is called monochromatic.

A 0-1 sequence \( t \) completely traverses \( s \) starting at vertex \( v \) if during traversal of \( s \) by \( t \) starting at \( v \) we visit all the vertices of \( s \). The 0-1 sequence \( t \) strongly traverses \( s \) starting at \( v \) if it completely traverses \( s \) starting at \( v \) and for every vertex there is a visit to that vertex which comes from left and continues to the right, or it comes from right and continues to the left i.e., the vertex is entered using one edge and left using the other one. A 0-1 sequence \( t \) is a (strong) universal traversal sequence for cycles of length \( n \) if for any cycle \( G \) of length \( n \), \( t \) completely (strongly) traverses \( G \) starting at any vertex.

Observe, any UTS for cycles of length \( n + 2 \) is a UTS for cycles of length \( n + 1 \) and a strong UTS for cycles of length \( n \). (Proofs of many claims are omitted in this extended abstract due to the space limitation; see \([K1]\) for proofs.)

3 An \( n^{O(\log n)} \) Universal Traversal Sequence

3.1 A Parity Contraction

Consider running a 1-run \( t \) of length \( l \) on a 1-run \( r_1 \) of length \( n \leq l \) starting at a left border vertex of \( r_1 \). Assume that \( r_1 \) is followed by a 0-run \( r_0 \). It is easy to see that if the parity of \( l \) and \( n \) are the same then we end up at the left border vertex of \( r_0 \), else we end up at the right border vertex of \( r_1 \). By symmetry, we get symmetric behavior for running a 0-run on a 0-run starting at the left border vertex, and for running a 1-run on a 0-run and a 0-run on a 1-run starting at the right border vertices. Hence, if the parity of the length of the traversed run corresponds to that of the traversal run we end up at the border vertex of the next run, otherwise we end up at the opposite border vertex of the same run.

This motivates the following definition. Let \( s \) be a non-monochromatic graph sequence, and let \( s \) be the same as \( r_0 r_1 \cdots r_{k-1} \), where every \( r_i \) is a monochromatic run and for every \( 0 \leq i < k \), \( r_i \) has an opposite color than \( r_{(i+1) \mod k} \).
For $0 \leq i < k$, let $r'_i$ be a run of the same color as $r_i$, and let the length of $r'_i$ be one if $r_i$ is of odd length and two otherwise. Then a parity contraction of $s$ is a sequence $r'_0 r'_1 \cdots r'_{k-1}$. I.e., we replace every run of odd length by one vertex run, and every run of even length by two vertex run.

There is a natural mapping of border vertices of original runs to border vertices of reduced runs. In the case of odd length runs, both border vertices are mapped to the same vertex. Observe that for traversing a parity contraction it suffices to consider (1,2)-sequences.

**Proposition 1.** Let $s$ be a non-monochromatic graph sequence of length $l$, and let $s_\oplus$ be its parity contraction. Let $v$ be a border vertex in $s$ and $v_\oplus$ the corresponding border vertex in $s_\oplus$. If a traversal (1,2)-sequence $t_\oplus$ strongly traverses $s_\oplus$ starting at $v_\oplus$ then $t$, obtained from $t_\oplus$ by inflating every run by $l + (l \mod 2)$ vertices, strongly traverses $s$ starting at $v$.

### 3.2 A Pair Contraction

Let $s$ be a (1,2)-sequence. We call a 0-run (1-run) of length 2 in $s$ a 00-pair (11-pair). A 0-run and a 1-run of length 1 are singletons and any maximal part of $s$ consisting of singletons is 01-run. (A 01-run does not necessary start by 0 and end by 1.) Let us consider a graph sequence $s$ containing a 01-run followed by a 11-pair and another 01-run: 01010 11 0101010101... Starting $t$ at the first vertex of $s$, we reach the left vertex $v_L$ of 11-pair after 5 digits of $t$. The next 1 in $t$ brings us to the right vertex $v_R$ of the 11-pair in $s$, the following 0 in $t$ takes us back to the vertex $v_L$, and so on. When 11-pair in $t$ comes we get to the vertex following the 11-pair in $s$ and we continue to run to the right. By symmetry, we observe similar effects for the other possible colorings of the pairs in $s$ and $t$.

Hence, if $s$ is a graph (1,2)-sequence and we traverse $s$ by a (1,2)-sequence $t$ then whenever we reach a pair in $s$ during that traversal we stay at that pair until a pair appears in $t$ and if the color of the pairs in $t$ and $s$ is the same then after that pair in $t$, the traversal goes to the right neighbor of the pair in $s$ otherwise to the left one. Even if several pairs are next to each other in $s$ not separated by any 01-run, this behavior of the traversal can be observed on each of these pairs.

This motivates the following construction. Let $s$ be a (1,2)-sequence containing at least one pair. A pair contraction $s_c$ of $s$ is a 0-1 sequence obtained from $s$ by first removing all singletons, and then replacing every pair by one vertex of the same color e.g., 01010 11 0101010101... is contracted to 110.... Note, the length of $s_c$ is at most half of the length of $s$.

**Proposition 2.** Let $s_c$ be a pair contraction of a (1,2)-sequence $s$ of length $l$. Let $t$ be a traversal sequence obtained from a strong traversal sequence $t_c$ of $s_c$ starting at $v_c$ by replacing every 1 by 011(01)$\lceil l/2 \rceil$, and every 0 by 001(01)$\lceil l/2 \rceil$. Then $t$ is a strong traversal sequence for $s$ starting at the left vertex of a pair corresponding to $v_c$, if $v_c$ is colored 0, and at the right vertex of that pair otherwise.
3.3 Putting Them Together

Propositions 1 and 2 give us a way to construct a universal traversal sequence of length $n^{O(\log n)}$. Given a strong universal traversal sequence $t_n$ for cycles of length $n$, we construct a strong UTS $t_{2n}$ for cycles of length $2n$ using the construction of Proposition 2 and then the construction of Proposition 1 as follows.

We first replace every 1 in $t_n$ by 011(01)$n$ and every 0 in $t_n$ by 001(01)$n$ as in Proposition 2 to get $t^{-3}_{2n}$. Then we prepend $(01)^n$ in front of $t^{-3}_{2n}$ to get $t^{-2}_{2n}$, which is a (1,2)-sequence. Now, we expand every run in $t^{-2}_{2n}$ by $2n = 2n + (2n \mod 2)$ vertices as in Proposition 1 to get $t^{-1}_{2n}$. Finally, we prepend $1^{2n}$ in front of $t^{-1}_{2n}$ to get $t_{2n}$. We call the prepended sequences $(01)^n$ and $1^{2n}$ loaders.

We claim that $t_{2n}$ is a strong UTS for cycles of length $2n$. Let $s$ be a graph sequence of length $2n$, let $s_\oplus$ be its parity contraction (assuming $s$ is not monochromatic), and let $s_c$ be the pair contraction of $s_\oplus$ (assuming $s_\oplus$ contains at least one pair.) Clearly, $|s_c| \leq n$, so $t_n$ strongly traverses $s_c$ starting at any vertex. The reader may easily verify using Proposition 1 that $t^{-2}_{2n}$ strongly traverses $s_\oplus$ starting at any vertex, and hence using Proposition 2 $t_{2n}$ strongly traverses $s$ starting at any vertex. (Loaders bring us to appropriate vertices so Propositions 1 and 2 are applicable.)

We need a strong UTS for cycles of length (let’s say) 5 to start this recursive construction of UTS’s.

4 An $O(n^c)$ Universal Traversal Sequence

A disadvantage of the previous construction is the super-polynomial length of the produced UTS. The reason for this length is a pessimism of the construction in two directions. The construction is pessimistic about the length of runs at each step, and about the number of rounds the construction has to be repeated.

A possible way how to reduce the overall length of the resulting UTS is to protect us against appearance of long runs. The way how we achieve this is by introducing two new stages in the construction which break down long runs. One of these stages takes place before parity contraction and the other one before pair contraction. The former stage reduces the length of 0-runs and 1-runs to constant length, the latter one reduces the length of 01-runs to constant length. A consequence of that run length reduction is an expansion of a traversal sequence in Propositions 1 and 2 just by a constant factor.

Both run breaking stages split long runs by inserting extra vertices in them. This actually increases the length of the graph sequence but in the consecutive contraction stages these increases are eliminated. The insertion is done in a way which ensures that any generated traversal sequence behaves on the new graph sequence and on the old one in the same way.

4.1 1-Run and 0-Run Breaking

This stage is aimed at breaking down long 0-runs and 1-runs. We are going to use the following simple proposition for $m$ equal 3 and 5.
Proposition 3. Let $m \geq 3$ be an odd integer. Then for any $l > m$ there is a $k$ and even $d$ such that $2 \leq d \leq 2m$ and $l = mk + d$.

Let $s$ be a graph sequence. We replace every run of length five and every run of length longer than six in $s$ by another sequence to obtain a graph $s'$ with all runs of short length. The replacement is done in the following way. Let $r$ be a 0-run or 1-run of length $l = 3k + d \in \{5, 7, 8, 9, \ldots\}$ in $s$, where $k$ and $d$ are from the previous proposition. If $r$ is a 0-run then we replace it by $r' = (000 01)^k 0^d$ otherwise we replace it by $r' = (111 10)^k 1^d$. We call this operation 0-run and 1-run breaking (or simply 1-run breaking.) We refer to $r$ as to the original sequence and to $r'$ as to the stuffed sequence.

Lemma 1. Let $s'$ be obtained from a non-monochromatic graph sequence $s$ by 1-run breaking. Let $v$ be a border vertex of some run in $s$ and let $v'$ be its corresponding vertex in $s'$. Let $(3,6)$-sequence $t$ strongly traverse $s'$ starting at $v'$. Then $t$, starting at $v$, completely traverses $s$, and all vertices in $s$, except possibly $v$, are strongly traversed.

This lemma is proven by focusing attention on behavior of (3,6)-sequences on original and stuffed runs. It can be shown that any (3,6)-sequence $t$ behaves the same on any original run $r$ and its corresponding stuffed sequence $r'$ in the following sense. If we run $t$ starting at leftmost (rightmost) input vertices $v$ and $v'$ of $r$ and $r'$, respectively, then $t$ keeps us in $r$ as long as it keeps us in $r'$, and it leaves $r$ and $r'$ at the same time on the corresponding output vertices, i.e., either at $v_l$ and $v'_l$, or $v_r$ and $v'_r$ (Fig. 1) While running $t$ we observe our position only after every whole run in $t$. (Sequence $t$ may temporarily leave $r$ or $r'$ during a run in $t$.) More formal treatment is given in the full version of this paper [K1].

![Fig. 1.](image-url)

Let $s$ and $s'$ be as in Lemma 1. Let $t$ be a traversal (3,6)-sequence which strongly traverses $s'$ starting at any vertex. Let sequence $t'$ be obtained from $t$ by prepending a run $w$ (a loader) of length $|s|$ of opposite color than the first run of $t$. Then Lemma 1 implies that $t'$ strongly traverses $s$ starting at any vertex.

We conclude that stuffing long runs of 0’s and 1’s is transparent for (3,6)-traversal sequences. The analysis is done explicitly for (3,6)-sequences but essentially identical analysis can be carried out for any $(2i+1, 4i+2)$-sequence, for $i \geq 1$, where we stuff long runs every $(2i+1)$ vertices.

The following proposition describes behavior of (3,6)-sequences on short runs.
Proposition 4. 1. Any (3,6)-sequence started at left border vertices or right border vertices behaves the same on runs 11, 1111 and 111111, where the output vertices are vertices of the other color adjacent to these runs in graph sequences that contain these runs. The same holds for 00, 0000 and 000000.

2. Any (3,6)-sequence started at left border vertices or right border vertices behaves the same on runs 1 and 111, where output vertices are as above. The same holds for 0 and 000.

Let $s'$ be a (1,2)-sequence and $t'$ be a strong traversal sequence of $s'$ starting at a vertex $v'$. We say that $t'$ extra strongly traverses $s'$ if every pair in $s'$ is traversed by $t'$ in such a way that it is entered from outside at one border vertex, walked back and forth on it and then left through the other border vertex.

Lemma 2. Let $s$ be a non-monochromatic graph sequence with every run of length from $\{1,2,3,4,6\}$. Let $s_{\oplus}$ be a parity contraction of $s$. Let $v$ be a border vertex in $s$, and $v_{\oplus}$ its corresponding vertex in $s_{\oplus}$. Let $t_{\oplus}$ be an extra strong traversal (1,2)-sequence of $s_{\oplus}$ starting at $v_{\oplus}$. Then $t$, obtained from $t_{\oplus}$ by replacing every run of length 1 by run of length 3, and every run of length 2 by run of length 6, completely traverses $s$ starting at $v$.

Note, if $s'$ is obtained from a non-monochromatic graph sequence $s$ by 1-run breaking, and $s_{\oplus}$ is obtained from $s'$ by parity contraction, then the length of $s_{\oplus}$ is at most the length of $s$.

Let $t_{\oplus}$ be a (1,2)-sequence which extra strongly traverses any graph (1,2)-sequence of length $|s|$ starting at any vertex. W.l.o.g. the first vertex in $t_{\oplus}$ has color 0. Let $t'$ be obtained from $t_{\oplus}$ by the construction from Lemma 2 and let $t$ be obtained from $t'$ by prepending a loader $1|s|$. Then $t$ strongly traverses $s$ starting at any vertex. Moreover, $t$ is a strong UTS for cycles of length $|s|$.

4.2 01-Run Breaking

This stage is aimed at breaking down long 01-runs. The method used here for breaking long 01-runs shares the same spirit with the previous method of breaking 0-runs and 1-runs. Let us describe this stage.

A $(5,10)$-01-sequence is a (1,2)-sequence in which every two consecutive pairs are separated by a 01-run and every 01-run is of length five or ten.

Let $s$ be a graph (1,2)-sequence containing at least one pair. We replace every 01-run in $s$ of length $l \in \{7,9,11,12,13,14, \ldots \}$ by a stuffed sequence $r'$ according to the following table. Let $l = 5k + d$ as in Proposition 3. Let $r_0 = 01010$, $r_1 = 10101$, $r'_0 = 01010$ 100100, and $r'_1 = 10101$ 011011.

$k$ even:

$$
(r_0r_1)^{k/2}(01)^{d/2} \Rightarrow (r'_0r'_1)^{k/2}(01)^{d/2}
$$

$$
(r_1r_0)^{k/2}(10)^{d/2} \Rightarrow (r'_1r'_0)^{k/2}(10)^{d/2}
$$

$k$ odd:

$$
(r_0r_1)^{(k-1)/2}r_0(10)^{d/2} \Rightarrow (r'_0r'_1)^{(k-1)/2}r'_0(10)^{d/2}
$$

$$
(r_1r_0)^{(k-1)/2}r_1(01)^{d/2} \Rightarrow (r'_1r'_0)^{(k-1)/2}r'_1(01)^{d/2}
$$
We can establish the following lemma which is an equivalent to Lemma [1]. The proof of Lemma [3] looks at 01-runs in $s$ and its corresponding stuffed sequences in $s'$, and shows that $(5,10)$-01-sequences behave the same on them with respect to vertices $v, v_l, v_r$ and $v', v'_l, v'_r$ (Fig. 2).

**Lemma 3** ($\approx$ [1]). Let $s'$ be obtained from a graph $(1,2)$-sequence $s$ containing at least one pair by 01-run breaking. Let $t$ be a traversal $(5,10)$-01-sequence ending with a 01-run of length ten. Let $v$ be a left vertex of some pair in $s$ if that pair has the same color as the first vertex of $t$, or let $v$ be a right vertex of that pair otherwise. Let $v'$ be a vertex corresponding to $v$ in $s'$. If $t$ extra strongly traverses $s'$ starting at $v'$ then $t$ extra strongly traverses $s$ starting at $v$.

![Fig. 2.](image)

Further, similarly to Proposition [4], any $(5,10)$-sequence beginning with 01-run and ending with a pair started at the leftmost vertices or rightmost vertices behaves the same on sequences 10, 1010, 101010, 10101010, and 1010101010, where the output vertices are vertices that are neighbors of the surrounding pairs in graph sequences containing these sequences (as in Fig. 2). The same holds for negated sequences, and for sequences 1, 101 and 10101, and their negations.

We conclude this section with the following lemma.

**Lemma 4.** Let $s$ be a $(1,2)$-sequence containing at least one pair with every 01-run having length in $\{1, 2, 3, 4, 5, 6, 8, 10\}$, and let $s_c$ be the pair contraction of $s$. Let $t_c$ be a strong traversal sequence of $s_c$ starting at the vertex $v_c$ corresponding to some pair in $s$. Let $v$ be the left vertex of that pair if the color of that pair is the same as the color of the first vertex in $t_c$, or let $v$ be the right vertex of that pair otherwise. Let $t$ be obtained from $t_c$ using the following rules:

<table>
<thead>
<tr>
<th>vertex replace by</th>
</tr>
</thead>
<tbody>
<tr>
<td>0 followed by 0</td>
</tr>
<tr>
<td>0 followed by 1 (or nothing)</td>
</tr>
<tr>
<td>1 followed by 1</td>
</tr>
<tr>
<td>1 followed by 0 (or nothing)</td>
</tr>
</tbody>
</table>

Then $t$ extra strongly traverses $s$ starting at $v$.

Note, if $s'$ is obtained from a $(1,2)$-sequence $s$ containing at least one pair by 01-run breaking, and $s_c$ is obtained from $s'$ by a pair contraction, then $|s| \geq 2|s_c|$.

Let $t_c$ be a UTS for cycles of length $\lceil |s|/2 \rceil$. W.l.o.g. the first vertex in $t_c$ has color 1. Let $t'$ be obtained from $t_c$ by the construction from Lemma [3] and let
Universal Traversal Sequences for Cycles of Length $O(n^{4.03})$

$t$ be obtained from $t'$ by prepending a loader $0(10)^{\lceil |s|/2 \rceil}$. Then $t$ extra strongly traverses $s$ starting at any vertex. Moreover, $t$ extra strongly traverses any graph $(1,2)$-sequence of length $|s|$.

We choose here to stuff 01-runs every 5 singletons. Similarly to the 1-run breaking stage, we could choose to stuff 01-runs every $2^i + 1$ singletons, for $i \geq 1$, and use for traversal a $(2i + 1, 4i + 2)$-01-sequence.

5 The Construction

Lemmas 4, 3, 2, and 1 combine to give a construction that converts a strong UTS for cycles of length $m$ into a strong UTS for cycles of length $2m$. This is presented explicitly in step 3 of the algorithm below. Here is the algorithm constructing a UTS for cycles of length $n$. Let $k = \lceil \log_2 n \rceil$.

The Algorithm:

1. Let $t_0$ be a strong UTS for cycles of length 6, that starts with color 1.
2. Apply the construction of Proposition 2 to $t_0$ to obtain a sequence $t'_1$, i.e., replace every 1 with 011(01)$^6$ and every 0 with 001(01)$^6$, and attach in front of $t'_1$ sequence $0(10)^6$ to get $t_1$.
3. For $i = 1, \ldots, k - 1$, construct a sequence $t'_{i+1}$ from $t_i$ by applying the following rules:

   replace by
   
   a 00-pair $[0010101]^5 00(10)^5$
   singleton 0 $[0010101]^2 00(10)^5$
   a 11-pair $[1101010]^5 11(01)^5$
   singleton 1 $[1101010]^2 11(01)^5$

   Prepend loader $0(10)^{6-4^i}[1101010]^{6-4^i} 11(01)^5$ in front of $t'_{i+1}$ to get a sequence $t_{i+1}$. (This step is a composition of the constructions from Lemmas 2 and 4.)
4. Apply the construction of Lemma 2 to $t_k$ to obtain a sequence $t'$, and attach sequence $1^n$ in front of $t'$ to get a sequence $t$.

Sequence $t$ produced by this algorithm is a universal traversal sequence for cycles of length $n$. By analyzing the algorithm it is easy to see that we can associate a derivation tree of logarithmic depth with the produced sequence. Every node in that tree corresponds either to a pair or a singleton, and certain nodes correspond to loaders. Hence, all the nodes, except the nodes corresponding to loaders, are of bounded degree. It is straightforward to traverse that tree in depth-first manner and to produce sequence $t$ using only logarithmic space. There is also a uniform $NC^1$ circuit that, given $1^n$, constructs that sequence.

The length of the produced sequence is proportional to the number of nodes at the last level of the derivation tree. The average number of descendants of a node corresponding to a pair or a singleton is at most 25.635. Thus, the number of nodes at $i$-th level of the tree satisfies the recurrence: $f(i + 1) \leq 25.635 f(i) +$
48 \cdot 4^i + 12. We set \( k = \lceil \log_2 n \rceil \) in the algorithm but it is actually enough to set \( k = \lceil \log_\sqrt{5} n \rceil \), because it can be shown that any graph sequence shrinks to the one-fifth of its original length during repeating the sequence of operations (1-run breaking, parity contraction, 01-run breaking, pair contraction) twice. Thus the length of the produced sequence is \( O(n^{\log_\sqrt{5} 25.635}) = O(n^{4.031}) \).

There may still be some room for slight improvement. If long 01-runs would be stuffed every seven singletons instead of five during the 01-run breaking and (7,14)-01-sequences would be used for traversal, the length of the produced sequence would possibly be \( O(n^{3.989}) \).

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References


On Universally Polynomial Context-Free Languages

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Abstract. A language is universally polynomial if its intersection with every NP-complete language is in P. Such a language would provide an automatic method for generating easy instances of intractable problems. In this note, we give a complete characterization of universally polynomial languages that are context-free, answering an open question in [4].

1 Introduction

It has been observed that most NP-complete languages have an infinite subset in P, i.e. most of these difficult languages have a nontrivial easy subset [2]. Whether this property holds in general is an open question with important consequences: a negative answer would refute the Isomorphism Conjecture for NP-complete sets, whereas a positive answer would imply that public-key cryptosystems based on NP-complete languages are inherently insecure.

Interestingly, it has also been observed that easy subsets for many NP-complete languages can be obtained by applying a common restriction to the problem domain. For example, 3-COLORING, CLIQUE, INDEPENDENT SETS, and PARTITION INTO CLIQUES are NP-complete problems that become polynomially solvable when the domain is restricted to chordal graphs [3]. This observation has led Demaine, Lopez-Ortiz, and Munro to ask whether there exists a universal restriction that would always yield an easy subset when applied to any NP-complete problem [4]; the existence of such sets would provide an automatic method for generating easy instances of intractable problems. Specifically, they introduced the concept of universally polynomial languages as those whose intersection with any NP-complete language is in P, and the weaker concept of universally simplifying languages as those whose intersection with any NP-complete language is not NP-complete. These two types of languages together are called universally easy languages. It turns out that C, the set of chordal graphs, is not a universally simplifying language, since the graph-theoretic problem of λ-COLORING remains NP-complete for chordal graphs [3]. Assuming that P ≠ NP, it follows that C is not universally polynomial.

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It is easy to see that all finite languages are universally polynomial. In practice, we are interested in nontrivial (i.e. infinite) universally polynomial languages that are efficiently computable. Note also that these concepts are interesting only if $P \neq NP$; otherwise every languages in NP would be universally polynomial and no language in NP would be universally simplifying. At present, a complete characterization of universally easy languages in P is not known. \cite{4} took a first step towards this goal by characterizing universally easy sets that belong to a natural subclass of P, namely the class of regular languages. Specifically, they showed that

1. all finite (regular) languages are universally polynomial;
2. all sparse (regular) languages are either universally simplifying or polynomial;
3. no nonsparse regular language is universally polynomial if $P \neq NP$.

In this paper, we give a complete characterization of universally polynomial languages that are context-free, answering a problem left open in \cite{4}. We show that the set of universally polynomial context-free languages consists of

1. all context-free languages if $P = NP$;
2. all sparse infinite context-free languages if $P \neq NP$ but $DEXT = NEXT$;
3. all finite languages if $DEXT \neq NEXT$.

We also obtain results for non-context-free languages. Specifically, we show that

1. no co-sparse language in $P$ is universally simplifying;
2. no co-sparse language in $NP$ is universally polynomial if $P \neq NP$.

The rest of this paper is organized into four sections. Relevant definitions and previously known results are stated in Section 2. The complete characterization of universally polynomial context-free languages appears Section 3 and the results on co-sparse universally easy languages appear in Section 4. Finally, some open questions are given in Section 5.

### 2 Preliminaries

All languages in this paper are over the binary alphabet $\Sigma = \{0, 1\}$. We use the terms language and set interchangeably. We assume that the reader is familiar with language classes in the Chomsky hierarchy (e.g. regular and context-free languages) and complexity classes (e.g. $P$, $NP$, $DEXT = \bigcup_c DTIME(2^{cn})$, and $NEXT = \bigcup_c NTIME(2^{cn})$). It is known that $REGULAR$ and $CONTEXT$-$FREE$ are subclasses of $P$. See \cite{6} for more details about these classes.

$A$ is a universally polynomial language if for every NP-complete language $C$, $A \cap C \in P$. Note that every finite languages is universally polynomial, and every language in NP is universally polynomial if $P = NP$.

$B$ is a universally simplifying language if for every NP-complete language $C$, $B \cap C$ is not NP-complete. Note that no language in NP is universally simplifying.
if $P = NP$. A language is universally easy if it is either universally simplifying or universally polynomial.

$L$ is a sparse language if the number of strings in $L$ of length at most $n$ is bounded above by a polynomial $p(n)$. $L$ is a tally language if $L \subseteq 0^*$. Note that all finite and tally languages are sparse. The existence of tally languages in NP – P has been characterized by the following theorem (due to Hartmanis) [1]:

**Theorem 1.** $\text{DEXT} \neq \text{NEXT}$ iff there is a tally set in NP – P iff there is a sparse set in NP – P.

$L$ is a bounded language if there are some binary strings $w_1, w_2, \ldots, w_n$ such that $L \subseteq w_1^*w_2^*\ldots w_n^*$. The following characterization of sparse context-free languages is well-known:

**Theorem 2 ([8,7]).** A context-free language is sparse iff it is bounded.

### 3 Characterization of Universally Polynomial Context-Free Languages

We begin by showing that $\Sigma^* \leq^p_P N$ for every nonsparse context-free language $N$, and hence none of them can be universally simplifying.

**Theorem 3.** No nonsparse context-free language is universally simplifying.

**Proof.** Let $L$ be a nonsparse context-free language over the binary alphabet $\Sigma = \{0,1\}$, and $G_L$ be a Chomsky normal form context-free grammar for $L$ without useless or nonreachable variables. Since $L$ is nonsparse, it is infinite. By the Pumping Lemma for context-free languages, there exists some variable $T$ that can make a derivation of the form:

$$T \Rightarrow^* X_0TY_0,$$

where $X_0$, $Y_0$ are binary strings, and $|X_0| > 0$ or $|Y_0| > 0$. Suppose in addition that the same variable $T$ can make another derivation of the form:

$$T \Rightarrow^* X_1TY_1,$$

where $X_1$ and $Y_1$ are binary strings, and $(X_0)^{|X_1|} \neq (X_1)^{|X_0|}$, or $(Y_0)^{|Y_1|} \neq (Y_1)^{|Y_0|}$.

Because $G_L$ contains no useless or nonreachable variables, by iterating the above two derivations, we can generate strings in $L$ of the form

$$ux_1x_2\ldots x_kvy_k\ldots y_2y_1w,$$

for every $k \geq 1$, where $u$, $v$, and $w$ are some binary strings, and each pair $(x_i, y_i)$ is either $((X_0)^{|X_1|}, (Y_0)^{|Y_1|})$, or $((X_1)^{|X_0|}, (Y_1)^{|Y_0|})$.

Define a polynomial-time, 1-1, invertible function $f$ from $\{0,1\}^*$ to $L$ as follows: given a binary string $z = b_1b_2\ldots b_n$ of length $n$,

$$f(z = b_1b_2\ldots b_n) = ux_1x_2\ldots x_nvy_1v\ldots y_2y_1w.$$
Since \( f \) is 1-1 and polynomially invertible, the image \( f(C) \) of any NP-complete set \( C \) under \( f \) is also NP-complete. This shows that \( L \) has an NP-complete subset and therefore is not universally simplifying.

So it remains to show that for nonsparse context-free languages \( L \), the additional derivation \( T \Rightarrow^* X_1TY_1 \) exists, where \( (X_0)^{|X_1|} \neq (X_1)^{|X_0|} \). We will show the contrapositive: if no such additional derivation exists, then \( L \) must be a bounded language and therefore must be sparse [8].

For each variable \( T \) of \( G_L \), consider all derivations of the form \( T \Rightarrow^* X_kTY_k \), where \( X_k \) and \( Y_k \) are binary strings (if they exist). If for any pair \((i, j)\), \((X_i)^{|X_j|} = (X_j)^{|X_i|} \) and \((Y_i)^{|Y_j|} = (Y_j)^{|Y_i|} \), then \( X_i \) and \( X_j \) can be written as \( X_i = (X_T)^{e_i} \) and \( X_j = (X_T)^{e_j} \), where \(|X_T| = \gcd(|X_i|, |X_j|)\). Similarly, \( Y_i \) and \( Y_j \) can be written as powers of some \( Y_T \) of length \( \gcd(|Y_i|, |Y_j|) \). Thus any derivation of \( T \) that begins with \( T \Rightarrow X_kTY_k \) will eventually generate a binary string of the form \((X_T)^*Z_T(Y_T)^*\). Here \( Z_T \) is a binary string, whose derivation tree contains no occurrences of \( T \) except at the root.

Now consider a derivation tree for any word in \( L \). For any long enough branch so that it has two occurrences of some variable \( T \), collapse the segment starting with the first and ending with the last occurrence of \( T \) and replace it with a node labeled by \((X_T)^*T(Y_T)^*\). Repeat this pruning process until no branch of the tree contains two or more occurrences of some variable. Each leaf of the pruned tree is associated with a nested regular expression of the form \((X_T)^*(X_T)^*\ldots(X_T)^*Z(Y_T)^*\ldots(Y_T)^*(Y_T)^*\), (which appear on the path from the root), and the variables \( T_i \) are all different. The number of such pruned trees as well as the number of labels that can be assigned to each leaf are constants depending only on \( L \). Hence every word in \( L \) has the form \( w_1^*w_2^*\ldots w_k^* \), where \( k \) is some constant depending only on \( L \). In other words, \( L \) is a bounded context-free language and hence it is sparse. \( \square \)

In view of Theorem 3, it is unlikely that any nonsparse context-free languages can be universally polynomial. In fact, either all or none of them are.

**Theorem 4.** The following are equivalent:

1. \( P = NP \);
2. every nonsparse context-free language is universally polynomial;
3. some nonsparse context-free language is universally polynomial.

**Proof.** \( \mathbf{1} \Rightarrow \mathbf{2} \): if \( P = NP \), then the intersection of any context-free language with an NP-complete language is a language in NP and hence must be in \( P \);

\( \mathbf{2} \Rightarrow \mathbf{3} \): obvious;

\( \mathbf{3} \Rightarrow \mathbf{1} \): suppose \( L \) is a nonsparse context-free language that is universally polynomial. The proof of Theorem 3 shows that \( L \) has a subset \( C \) which is NP-complete. Hence \( L \cap C = C \) is both NP-complete and in \( P \), so \( P = NP \). \( \square \)

We now turn our attention to sparse context-free languages. Using the same technique, we show that \( 0^* \leq^p_1 L \) for every infinite sparse context-free language \( L \), and hence none of them are likely to be universally polynomial. In fact, either none or all of them are. Theorem 1 is crucial in the proof of this characterization.
**Theorem 5.** The following are equivalent:

1. $\text{DEXT} = \text{NEXT}$;
2. every infinite sparse context-free language is universally polynomial;
3. some infinite sparse context-free language is universally polynomial.

**Proof.** (1) $\Rightarrow$ (2): iff $\text{DEXT} = \text{NEXT}$, then the intersection of any infinite sparse context-free language with an NP-complete set is a sparse set in NP and therefore must be in P;

(2) $\Rightarrow$ (3): obvious;

(3) $\Rightarrow$ (1): let $L$ be an infinite sparse context-free language. It follows from the Pumping Lemma for context-free languages that there exist strings $x, y, u, v, w$ such that $|y| + |v| > 0$ and $L' = \{xy^iuv^iw : i \geq 0\}$ is an infinite subset of $L$. The map $f(0^i) = xy^iuv^iw$ is a polynomial-time isomorphism between $0^\ast$ and $L'$ such that for any tally language $T$, $T$ is in P iff $f(T)$ is also in P.

Given a tally language $T$ in NP, we construct an NP-complete language $C_T$ such that $C_T \cap L = f(T)$ as follows. Define $S = \{wx : |w| = |x| \wedge x \in \text{SAT} \wedge wx \notin L\}$. $S$ is clearly in NP. Furthermore, SAT $\leq^p_m$ $S$ via the following reduction: on input $x$, enumerate strings $w$ of length $|x|$ until $wx \notin L$ is satisfied, and output $wx$. Since $L$ is sparse, such a string can be found in polynomial time. Furthermore, $x \in \text{SAT}$ iff $wx \in S$. Thus $S$ is NP-complete. Note also that $S \cap L = \emptyset$. Now define $C_T = S \cup f(T)$. Clearly $C_T$ is also NP-complete and $C_T \cap f(T) = f(T)$.

Since $L$ is universally polynomial, $f(T)$ is in P for all tally set $T$ in NP. Hence all tally sets in NP are in P, so $\text{DEXT} = \text{NEXT}$. \hfill \Box

Combining Theorems 4 and 5 we can now describe precisely the class of universally polynomial context-free languages provided we know the relationships between P vs. NP and DEXT vs. NEXT. Note that DEXT $\neq$ NEXT implies that P $\neq$ NP.

**Corollary 1.** The class of universally polynomial context-free languages consists of

1. all context-free languages if $P = \text{NP}$;
2. sparse context-free languages if $P \neq \text{NP}$ but DEXT = NEXT;
3. finite languages if DEXT $\neq$ NEXT.

## 4 Universally Easy Co-sparse Languages

We extend the results obtained in Section 3 for non-context-free languages whose complements are sparse. Again, it is unlikely that any of the co-sparse languages in NP is universally polynomial, but if one is, then all of them are.

**Theorem 6.** No co-sparse language in P is universally simplifying.
Proof. Let \( L \in P \) be a language whose complement is sparse. Define a polynomial-time 1-1 function \( f : \Sigma^* \mapsto L \) as follows: on input \( w \), \( f \) searches for the lexicographically first string \( v \) of length \( |w| \), such that \( vw \in L \) and outputs \( vw \). Such a string exists and can be found in polynomial time since \( L \) is in \( P \) and \( \overline{L} \) is sparse. Clearly \( f \) is one-one.

For any NP-complete language \( C \), its image \( f(C) \subseteq L \) is also in NP, since \( f \) is polynomially honest, and therefore NP-complete. Hence \( L \) is not universally simplifying.

Theorem 7. If \( P \neq NP \), then no co-sparse language is universally polynomial.

Proof. Suppose \( P \neq NP \). Let \( C \) be an NP-complete language and define \( C' = \{xw : |x| = |w| \land w \in C\} \). \( C' \) is also NP-complete. If there is a co-sparse language \( X \) such that \( X \cap C' \in P \), then we show that \( C \in NP \), a contradiction.

Let \( |X| \leq n^k \) be bounded by some polynomial \( n^k \). The following is a polynomial-time algorithm for \( C \): on input \( w \), enumerate the first \( |2w|^k + 1 < 2^{|w|} \) strings of length \( |w| \), and for each such string \( x \), check whether \( xw \in X \cap C' \). It is clear that \( w \in C \) iff one such test succeeds.

Corollary 2. The following are equivalent:
1. \( P = NP \);
2. every co-sparse language in NP is universally polynomial;
3. some co-sparse language in NP is universally polynomial.

5 Conclusions

We have characterized the class of universally polynomial context-free languages. Under the assumption that i) \( P \neq NP \) but ii) \( DEXT = NEXT \), an infinite universally polynomial language in \( P \) exists. It is not known whether the second assumption is necessary for the existence of such sets. It would also be interesting to investigate the relationship between universally easy languages and NP-complete languages without an infinite P-subset.

References

Separating Oblivious and Non-oblivious BPs

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Abstract. This paper gives an exponential separation on the depth of branching programs (BPs) between oblivious and non-oblivious BPs. Namely, there is a difference just like the difference between sequential and NC computation: (i) There is a Boolean function $f_1$ of $N$ variables which can be computed by a polynomial-size, syntactic BP with a depth of $2\log N - \log \log N + 1$ but cannot be computed by any oblivious BPs with a depth of $(2 - \varepsilon)N$ for some $\varepsilon \in o(1)$. (ii) Similarly, there is an $f_2$ computed by a syntactic depth of $\log^3 N$ but not by an oblivious depth of $\Omega(\log N \log N)$. (iii) We also show that any (unrestricted) BP of depth $t$ can be simulated by an oblivious BP with a depth of $N + [(t - \log N)/(\log \log N + C)] \cdot N$. The third result implies that $f_1$ cannot be computed by any BP with a depth less than $\log N + \log \log N$ and $f_2$ not with a depth of $o(\log N \cdot \log \log N)$. Note that most bounds in this paper include constant factors and lower-degree terms.

1 Introduction

One of the primary goals of theoretical computer science is to prove non-linear lower bounds for the size of Boolean circuits. However, in spite of serious efforts for several decades, we are still at a very elementary stage and are struggling for good research hints. Probably for this reason, several researches have been shifting their attention from Boolean circuits to Boolean branching programs (BPs): (i) BPs are also a general computation-model whose size and depth are naturally related to space and time [8]. (ii) It is relatively easy to consider reasonable restrictions which provide interesting subclasses, such as oblivious models discussed in this paper, for more attainable lower-bound arguments.

Not surprisingly, therefore, we have a large literature. First of all, Ajtai [2] recently made a big breakthrough by showing a nonlinear depth lower bound for unrestricted BPs (no nontrivial lower bounds for even read-twice BPs were known before.). In [13], Thathachar gave an exponential separation on the size between $k$-sBPs (read-$k$ syntactic BPs) and $(k + 1)$-sBPs for all $k$, extending the previous result [7] that separates only between $k = 1$ and $k = 2$. Then, Beame, Saks and Thathachar [10] obtained the first (again exponential on the size) separation result between semantic and syntactic BPs, i.e., between $k$-BPs.

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Separating Oblivious and Non-oblivious BPs 29

and \(k\)-sBPs, by giving a Boolean function \(f_k\) which can be computed by a 2-BP of polynomial size but cannot be computed by any \(k\)-sBP of polynomial size. For oBPs (oblivious BPs) and OBDDs, similar separation results have been known for longer time, including those between \(k\)-oBPs and \((k+1)\)-oBPs [13], between \(k\)-OBDDs and \((k+1)\)-OBDDs [9,3] and between \(2\)-oBPs and \(\bigcup_k k\)-OBDDs [11]. (The separation between \(k\)-oBPs and \((k+1)\)-oBPs is mentioned only implicitly in [13].)

In this paper, we focus our attention not on the size but on the depth of BPs which is an equivalently popular measure (e.g., [4,6]). Interestingly, our results have exactly the same flavor as typical results on parallel vs. sequential computation. Remember what we have usually done in the research of parallel computation: (i) For a given problem (or a Boolean function) \(Q\), we design an efficient (poly-log depth) parallel algorithm \(A\) and then (ii) claim its efficiency by comparing the time complexity of the standard sequential simulation of \(A\) and the time complexity of the best known sequential algorithm for \(Q\). This paper just follows this approach: (i) There is a Boolean function \(f_1\) of \(N\) variables which can be computed by a syntactic BP of depth \(2 \log N - \log \log N + 1\) but cannot be computed by any oblivious BPs of depth \((2 - \varepsilon)N\) for some \(\varepsilon \in o(1)\).

Also, (ii) there is an \(f_2\) which is computed by a syntactic depth of \(\log^2 N\) but not by an oblivious depth of \(\Omega(N \log N)\). Thus, a poly-logarithmic depth is enough for some functions which need a polynomial (i.e., exponentially larger) oblivious depth. Our third result is on the simulation; (iii) any BP of depth \(t\) can be simulated by an oblivious BP with a depth of \(N + \left\lceil \frac{(t-\log N)}{(\log \log N + C)} \right\rceil \cdot N\).

This simulation result implies that \(f_1\) cannot be computed by any BP with a depth of less than \(\log N + \log \log N\) and \(f_2\) not with a depth of \(o(\log N \cdot \log \log N)\). Thus those lower bounds are somewhat close to the upper bounds in (i) and (ii) above, but are not tight. It should be noted that our results in this paper give specific constant factors and even lower-degree terms.

As for lower-bounds on the depth of oblivious BPs, three Boolean functions have appeared in the literature: two have an \(\Omega(N \log N)\) lower bound [4] and the third one has an \(\Omega(N \log^2 N)\) lower bound [6]. Our second function \(f_2\) is a modification of one of the two functions in [4]. Similar modification is also possible to the function in [6] and we can design an efficient syntactic BP. Unfortunately, we can only get an \(O(N^{\varepsilon})\) upper bound for this syntactic depth at this moment (see Section 5).

It is hard to cover all the related research in this detailed abstract. Among others, many authors have discussed BPs which can read most variables only once but can read the remaining few variables many times [e.g., [12]]. OBDDs have a long history of research which includes a lot of papers discussing their practical merits for computer-aided design and hardware verification [e.g., [5]].

In the rest of this paper, we first give definitions and notations in the next section. Then we give the simulation result in Section 3, which will help understanding basic differences between oblivious and non-oblivious BPs. Sections 4 is for the two separation results and finally we briefly mention some related results in Section 5.
2 Definitions

Fig. 1 illustrates a simple example of a BP. A BP is an acyclic directed graph with a single source vertex of zero in-degree. Each sink vertex of zero out-degree is labeled either 0 or 1. Each non-sink vertex is labeled some variable $x_i$ among $x_1, x_2, \ldots, x_N$ and has two outgoing edges labeled 0 and 1. Once a 0/1-assignment to the variables $x_1$ through $x_N$ is given, we can determine a unique computation path from the source vertex to some sink vertex labeled 0 or 1. For example, if $(x_1, x_2, x_3) = (1, 0, 1)$, the unique path ends with the sink vertex labeled 1 in Fig. 1. Thus a BP computes a Boolean function of $N$ variables. The size of a BP is the number of non-sink vertices and the depth of a BP is the number of non-sink vertices on the longest computation path. In this paper, all the BPs have a size bounded by some polynomial in $N$, so we often omit mentioning it.

![Fig. 1. example of a BP](image)

As shown in the example in Fig. 1, each computation path can read the variables in a different order. If all the computation paths read the variables in the same order, such a BP is called an oblivious BP. (Unrestricted or semantic) BPs and syntactic BPs have no such restriction on the order of reading variables. Instead there is an important difference between them in the way of measuring the depth. See Fig. 1 again: If one looks at this BP carefully, it turns out that the path $x_1 x_2 x_1 x_3 x_1$ is not actually followed by the BP since it would mean that the value of $x_1$ is 0 when it is read for the first time and is 1 for the second time. Such a computation path is said to be not realizable. Syntactic BPs take into account all computation paths whether or not they are realizable and its depth is the length of the longest path among them. Semantic BPs (or simply BPs) take into account only realizable computation paths and its depth is the length of the longest, realizable path.
A BP is said to be \textit{tree-like} if each vertex (other than the source) has only one in-degree. Any BP can be changed to a tree-like BP simply by splitting a vertex of in-degree \(k\) into \(k\) vertices of in-degree one from top to bottom. Note that the depth does not change but the size can enlarge greatly. Obviously, any tree-like BP has an equivalent tree-like BP such that no variable appears twice or more in every path.

3 Simulation by Oblivious BPs

Recall that the relationship between BPs and oblivious BPs is somewhat similar to the relationship between parallel and sequential computation.

\textbf{Theorem 1.} Suppose that a Boolean function \(f\) of \(N\) variables is computed by a BP of depth \(t\). Then \(f\) can be computed by an oblivious BP of depth 
\[ N + \left\lceil \frac{(t - \log N)}{\log \log N + C} \right\rceil \cdot N. \]

\textbf{Proof.} For a given BP \(Q\) of depth \(t\) we construct an oblivious BP \(P\). The construction is basically to divide \(Q\) into blocks of depth \(\log \log N\) and simulate each of them by oblivious BPs of depth \(N\). However, a special treatment can be done for the top \(\log N\)-depth portion of \(Q\) (see Fig. 2).

\begin{figure}[h]
\centering
\includegraphics[width=\textwidth]{oblivious_simulation.pdf}
\caption{Simulation by Oblivious BPs}
\end{figure}
**Lemma 1.** Any BP of depth $\log N$ can be simulated by an oblivious BP of depth $N$.

**Proof.** We first change the given BP, say, $M_1$, into a tree-like BP, $M_2$, which includes $N-1$ vertices or at most $N-1$ variables. Suppose that $M_2$ looks like Fig. 3(a). Then an equivalent oblivious BP can be constructed simply by “inserting missing vertices” in each path of $M_2$ as shown in Fig. 3(b). For example, we insert $v_2, v_4, v_5$ and $v_7$ into the path of $v_1v_3v_6$ of $M_2$. Each inserted vertex has only one out-degree and therefore the size does not blow up. 

It should be noted that this construction is not always possible if the variable order in the target oblivious-BP is restricted. For example, if the variable order of the oblivious BP is the opposite of the above width-first order, then the size turns out to be exponential. This is the reason why we cannot divide $Q$ into all $\log N$-depth blocks.

**Lemma 2.** Let $C > 0$ be any constant and $S = v_1v_2\cdots v_N$ be any variable order. The any BP of depth $\log \log N + C$ whose variable set is a subset of $\{v_1, v_2, \cdots, v_N\}$ can be simulated by an oblivious BP of depth $N$ whose variable order is $S$.

![Fig. 3. Basic idea of simulation](image_url)

**Proof.** Since the given BP, say, $M_1$, is of depth $\log \log N + C$, the number of vertices is $O(\log N)$, namely $M_1$ looks at only $O(\log N)$ variables. Hence we can immediately make an oblivious BP, $M_2$, whose shape is a completely binary tree of size $2^{O(\log N)}$ and whose variable order does not contradict to $S$. Finally we can just insert missing variables as before.

Now we divide the original BP $Q$ into blocks as shown in Fig. 2. The first level includes vertices from the first level until the $(\log N + 1)$th level, the second block includes vertices from the $(\log N + 1)$th level until the $(\log N + \log \log N + C + 1)$th level. (Namely, vertices at the $(\log N + 1)$th level are shared by the first and the second blocks.) In general the $k$th ($k \geq 3$) block starts at $(\log N + (k - 1)(\log \log N + C) + 1)$th level and ends at $(\log N + k(\log \log N + C) + 1)$th level.
Thus $Q$ is divided into one block of depth $\log N$ (vertices at the lowest level at each block are viewed as sink nodes and are excluded in counting the depth), and $\lceil (t - \log N)/(\log \log N + C) \rceil$ blocks of depth $\log \log N + C$, excepting the last block which may be shorter.

Each block can be viewed as a set of (a polynomial number of) BPs. Each of them is transformed into a tree-like BP but we do not change vertices at the lowest level, i.e., they can accept an arbitrary number of incoming edges. Now these BPs are transformed into oblivious BPs as follows:

1. The first block is transformed into an oblivious BP using the construction in Lemma 1.
2. For each BP in other blocks, we use the construction in Lemma 2, where the variable sequence $S$ for all these BPs must be the same but may be any (fixed) one.

Note that we did not change the vertices which are shared by neighboring two blocks (whose labels may have been changed) and therefore those transformed BPs are naturally combined into a single BP which is equivalent to $Q$. One can see that this BP is oblivious and its depth is $\log N + \lceil (t - \log N)/(\log \log N + C) \rceil \cdot N$.

\[ \square \]

4 Separation Results

In this section $\log N$ and $\log \log N$ always denote $\lceil \log N \rceil$ and $\lceil \log \log N \rceil$, respectively. We introduce two Boolean functions which give an exponential separation on the depth between oblivious and non-oblivious BPs.

**Theorem 2.** Let $\varepsilon$ be any function such that $\varepsilon > 2 \frac{\log \log N}{\log N}$ for infinitely many integers $N \geq 0$ (e.g., $\varepsilon$ may be $\frac{(\log \log N)^{1+\delta}}{\log N} \in o(1)$). Then there exists a Boolean function, $f_1$, of $N$ variables such that $f_1$ can be computed by a syntactic BP of depth $2 \log N - \log \log N + 1$ but cannot be computed by any oblivious BP of polynomial size whose depth is $(2 - \varepsilon)N$.

**Proof.** A Boolean function $f_1(x_1, x_2, \ldots, x_N)$ is defined as follows: Its variables $x_1, x_2, \ldots, x_N$ are partitioned into blocks $\beta_0, \beta_1, \ldots, \beta_{\lceil \frac{N}{\log N} \rceil - 1}$. The first block $\beta_0$ consists of the first $\log N - \log \log N$ variables, i.e., $x_1$ through $x_{\log N - \log \log N}$, and each of subsequent $\beta_i$’s ($i \geq 1$) consists of subsequent $\log N$ variables (the last one may be shorter again). To decide the value of $f_1$: (i) We first look at block $\beta_0$ and regard its value as a binary number $b_0$. (ii) Then we go to block $\beta_{b_0+1}$ and again regard its value as a binary number $b_1$. (iii) The value of $f_1$ is equal to the value (0 or 1) of $x_{b_1}$.

**Upper Bound.** It is not hard to see that $f_1$ can be computed by a tree-like BP $Q$ of depth $2 \log N - \log \log N + 1$. Its size is obviously bounded by a polynomial because its depth is logarithmic. $Q$ first reads variables in $\beta_0$ which includes $\log N - \log \log N$ variables and due to its value $Q$ decides which $\beta_i$ of $\log N$ variables should then be read. Finally $Q$ reads $x_{b_1}$ and goes to a sink vertex of 0 or 1.

**Lower Bound.** Suppose that there is an oblivious BP, $M$, of depth $(2 - \varepsilon)N$ whose variable sequence is $S$. For simplicity we assume that the length of $S$ is
even and let \( S_1 \) and \( S_2 \) be the first and the second halves of \( S \). The following simple lemma is important:

**Lemma 3.** There are at least \( \frac{\varepsilon}{2} N \) variables which appear only in \( S_1 \) and also at least other \( \frac{\varepsilon}{2} N \) variables which appear only in \( S_2 \).

**Proof.** Since \( |S_1| = (1 - \frac{\varepsilon}{2})N \), there are at least \( \frac{\varepsilon}{2} N \) variables which do not appear in \( S_1 \). Those variables must appear in \( S_2 \) since \( f_1 \) clearly depends on all the variables. Similarly for \( S_2 \). 

Let us call variables only appearing in \( S_1 \) A-variables and variables only in \( S_2 \) Z-variables. Recall that \( |\beta_0| = \log N - \log \log N \) and there are at least \( \frac{\varepsilon}{2} N \) Z-variables. Therefore, some block, \( \beta_i(i \geq 1) \), contains at least

\[
\frac{\varepsilon/2}{} \frac{N - (\log N - \log \log N)}{\log N^1} = \frac{\varepsilon}{2} \log N - O(\frac{\log^2 N}{N})
\]

Z-variables. Although details are omitted, we can show that the second term \( O(\frac{\log^2 N}{N}) \) is so small and does not cause any problem if we omit it. In the following we do not write it just for simplicity. Select arbitrary \( \frac{\varepsilon}{2} \log N \) Z-variables in \( \beta_i \) and let \( BZ_i \) be the set of those variables. Also let \( B\overline{Z}_i \) be the set of the remaining \( (1 - \frac{\varepsilon}{2}) \log N \) variables in \( \beta_i \). Recall that the value assigned to \( \beta_i \)-variables is a pointer to some variable \( x_i(1 \leq i \leq N) \). Hence we can define the following equivalence classes on the set of the whole \( N \) variables:

\[
E_j = \{ x_i \mid \text{the pointer to } x_i \text{ (= the value } i) \}
\]

has the same 0/1 assignments (= j) to \( B\overline{Z}_i \) variables}

Note that the number of these equivalence classes is \( 2(1 - \frac{\varepsilon}{2}) \log N = N^{1 - \frac{\varepsilon}{2}} \), and therefore at least one \( E_j \) contains \( \frac{\varepsilon}{N^{1 - \frac{\varepsilon}{2}}} = \frac{\varepsilon}{2} \cdot \frac{N^2}{N} \) A-variables or more.

Now recall that all those A-variables appear in \( S_1 \) which is the first half of the variable sequence, and any of them, \( x_k \), can be selected by assigning some proper values to \( B\overline{Z}_i \) variables. (Reason: We first Fix the \( \beta_0 \) variables to show the \( \beta_i \) block. Then fix \( B\overline{Z}_i \) variables so that it will show the \( E_j \) equivalence class which includes \( x_k \). Now we finally fix \( B\overline{Z}_i \) variables to show \( x_k \).) Since those A-variables appear in \( S_1 \) and all the \( B\overline{Z}_i \) variables appear in \( S_2 \), the BP \( M \) has to remember all the values of the A-variables at the end of \( S_1 \). This needs \( 2^{\frac{1}{2} \cdot N^{\frac{\varepsilon}{2}}} \) vertices there, which is not bounded by any polynomial if \( \varepsilon \) satisfies the condition of the lemma. 

**Theorem 3.** There exists a Boolean function, \( f_2 \), of \( N \) variables such that \( f_2 \) can be computed by a syntactic, read-once BP of depth \( \log^3 N \) but needs an \( \Omega(N \log N) \) depth to be computed by oblivious BPs.

**Proof.** Before giving the definition of \( f_2 \), we shall review the lower-bound result in [4].

**Lemma 4** [4]. There is a Boolean function, \( f_2^* \), of \( N \) variables for which any oblivious BP need a depth of \( \Omega(N \log N) \).

Boolean function \( f_2^* \) looks like this: The variables are \( x_1, \ldots, x_n, y_1, \ldots, y_n \) and each variable takes value 0,1 or 2 i.e., they are 3-way variables. Actually, we use two Boolean variables for each \( x_i \) and \( y_i \), and hence \( N = 4n \). It is defined that
Separating Oblivious and Non-oblivious BPs

$f_2^* = 1$ iff the 0/1-sequence of $x_1, \cdots, x_n$ obtained by neglecting the variables having value 2 is equal to the 0/1-sequence of $y_1, \cdots, y_n$. The basic idea of their proof is as follows: Suppose that an oblivious BP, $M$, computes $f_2^*$ whose variable sequence is $S$ and that most variables are read at least $k$ times in $S$. Then it turns out, by a Ramsey-like theorem, that we can select $(\frac{1}{2})^{ck} n$ $x_i$-variables and the same number of $y_i$-variables for some constant $c > 0$ such that $S$ is divided into $ck$ blocks each of which contains only selected $x_i$-variables or only selected $y_i$-variables. Since all the other variables can be killed by assigning 2, a simple counting argument gives us

$$W^{ck-1} \geq 2^{n(\frac{1}{2})^{ck}},$$

where $W$ is the maximum width at each boundary of blocks. For $W$ to be bounded by a polynomial, we need that $k = \Omega(\log n)$, which implies the lower bound of the lemma.

This argument is perfectly good for their result but a careful observation gives us the following improvement:

**Observation.** We only need less survived variables than $n(\frac{1}{2})^{ck}$ ones to obtain a superpolynomial bound. Namely $\log^2 n$ ones are enough since in order to satisfy

$$W^{ck-1} \geq 2^{|\log^2 n|},$$

$W$ needs to go beyond a polynomial if $k = o(\log n)$.

This observation seldom improves their result (only a constant factor), but it is very important in this paper. We again use a similar addressing scheme as before to show these surviving variables.

**Definition of the Function.** $f_2(X_1, X_2)$ is defined as follows: Both $X_1$ and $X_2$ consist of $n \log N' + N'$ variables like

$$X_i = Q_{i,1}, \cdots, Q_{i,n}, x_{i,1}, \cdots, x_{i,N'}, \quad (i = 1, 2),$$

where $n = \log^2 N'$ and each $Q_{i,j}$ includes $\log N'$ variables. Note that the whole number $N$ of variables is $N = 2(n \log N' + N')$. When 0/1-values are assigned to the variables, let $s_{i,j}$ ($i = 1, 2$) be the sum of the values (the binary numbers as before) for $Q_{i,1}, \cdots, Q_{i,j}$. Then, $f_2 = 1$ if and only if (i) the value for $Q_{i,j} > 0$ for all $i$ and $j$, (ii) $s_{i,n} \leq N'$ for both $i = 1$ and 2, and (iii) $x_{1,s_{1,j}} = x_{2,s_{2,j}}$ for all $1 \leq j \leq n$. Namely the 0/1-sequence of $x_{1,j}$’s which are selected by $Q_{1,1}, \cdots, Q_{1,n}$ must be the same as the 0/1-sequence of similarly selected $x_{2,j}$’s.

**Lower Bound.** We omit the lower-bound proof since it is very similar to Lemma 4.

**Upper Bound.** Since the previous BP given in the proof of Theorem 2 had a logarithmic depth, it was able to be constructed as a tree-like BP. This time, however, tree-like BPs obviously blow up and we need a careful description for DAG-like BPs. To this end, we introduce what we call a “RAM-model” $M$ for BPs: (i) $M$ has a level-counter $i$ and an internal state $S$ of $O(\log N)$ bits. (ii) Its one-step move is determined by the values of $i$ and $S$, which includes reading some variable, and changing the values of $i$ and $S$ due to the current values of $i$
and $S$ and the value of the variable just read. (iii) $M$ may read more than one variable in its one-step move (multiple read).

It is not hard to see that a program written by this rule can be immediately transformed to a BP by changing all different states into different vertices. Its size is bounded by a polynomial since the internal state has only $O(\log N)$ bits. Its depth is the maximum level, but if multiple read is used then we have to increase it appropriately.

Now we give our RAM-model program $M$ which computes $f_2$. $M$ reads variables in the following order:

$$Q_{1,1}, x_{1,s_{1,1}}, Q_{2,1}, x_{2,s_{2,1}}, Q_{1,2}, x_{1,s_{1,2}}, \ldots, Q_{1,n}, x_{1,s_{1,n}}, Q_{2,n}, x_{2,s_{2,n}}$$

Here which variable $x_{i,s_{i,j}}$ is depends on the value of $Q_{i,j}$. Also, all variables in $Q_{i,j}$ are read in one step, which will not cause any problem since $|Q_{i,j}| = \log N' \leq \log N$. Internal state of $M$ has the form of $S = (P_1, P_2, z)$, where $P_1$ holds the position of variable $x_{1,j}$ which was read most recently, i.e., that is the value of $s_{1,j}$ at that moment. $z$ needs only one bit to hold the value of $x_{1,s_{1,j}}$ until it is compared with $x_{2,s_{2,j}}$. Thus, the size of $S$ is $2 \log N' + 1$ or $O(\log N)$. The level counter $i$ changes as follows:

$$(1, 1)Q(1, 1)x(2, 1)Q(2, 1)x(1, 2)Q(1, 2)x \cdots (2, n)Q(2, n)x,$$

where $(i, j)_Q$ means that it is the time when $M$ is to read $Q_{i,j}$ and $(i, j)_x$ means that it is the time when $M$ is to read $x_{i,s_{i,j}}$. Note that its initial state is $(0, 0, 0)$.

Now we describe a one-step behavior at each level.

(i) The current level is $(i, j)_Q$. Read $Q_{i,j}$. Suppose that the current state is $(P_1, P_2, z)$. Then if $Q_{i,j} = 0$ or $P_i + Q_{i,j} \geq N'$, then go to $[0]$ (= a sync-node labeled by 0). Otherwise, change its state to $(P_1 + Q_{1,j}, P_2, z)$ when $i = 1$ and to $(P_1, P_2 + Q_{2,j}, z)$ when $i = 2$ and go to the next level.

(ii) The current level is $(1, j)_x$. Read $x_{1,P_1}$, change its state to $(P_1, P_2, x_{1,P_1})$, and go to the next level.

(iii) The current level is $(2, j)_x$. Read $x_{2,P_2}$. If the current $z$ is equal to $x_{2,P_2}$ then if $j = n$ then go to $[1]$ else just go to the next level. If $z \neq x_{2,P_2}$ then go to $[0]$.

One can see that this program reads at most $2n$ $Q_{i,j}$’s and $2n$ $x_{i,j}$’s and hence the depth of the equivalent BP is $2n \log N' + 2n = O(\log^3 N)$. (Recall that $n = \log^2 N'$ and $N = (n \log N' + N')$. Also this BP never reads a single variable twice, namely, it is a read-once BP.

\[\square\]

## 5 Some Remarks

Here are a few remarks which might help:

**Remark 1.** One might have noticed that there is a trivial function of $N$ variables which is computed by a BP of depth $\log N$ but needs an $N$ oblivious depth. One such example is the function which is computed by the BP that has a form of a complete binary tree and all the vertices read different variables.
The final value is the value of the variable which is read at the end of each path. Clearly this function depends on all the variables and therefore any equivalent oblivious BP needs the depth which is equal to the number of variables. Such a separation is optimal due to Theorem 1. Our real purpose in this paper is not to give such a single example (which is enough only to prove a separation) but to give more general evidence that oblivious BPs need exponentially more depth than their non-oblivious counterparts for a wide range of Boolean functions.

**Remark 2.** By a small modification of $f_2$ of Theorem 3, we can obtain a function $f'_2$ which needs a non-linear oblivious depth but can be computed by a BP of depth $\log^{2+\varepsilon} N$ for any constant $\varepsilon > 0$. The idea is that if $k$ is a constant, then the width $W$ which satisfies $W^{ck-1} \geq 2(\log n)^{1+\varepsilon}$ is not bounded by any polynomial. Thus the construction of $f'_2$ is exactly the same as $f_2$ except for $n = \log^{1+\varepsilon} N'$.

**Remark 3.** The best depth lower bound for oblivious BPs is $\Omega(N \log^2 N)$ in [6]. The Boolean function used there also uses 3-way variables, i.e., variables having value 2 are all “killed.” Therefore, we can also build our addressing scheme into this function. Unfortunately, their lower-bound proof needs a polynomially many variables to survive. Thus this construction still needs $O(N^\varepsilon)$ depth to be computed by unrestricted BPs.

### 6 Conclusion

An apparently interesting attempt is to try to find a nontrivial separation between oblivious and non-oblivious depth which is optimal. As mentioned before, our Theorems 2 and 3 are not optimal with respect to Theorem 1. Theorem 3 is a little far from the optimal value but Theorem 2 is much closer. One possibility is to reduce the size of $\beta_i$ ($i \geq 1$) from $\log N$ to, say, $(\log \log N)^k$. If we do so, then it can no longer make pointers to all the variables. However, one can see that it would be enough to have pointers to some subset of the variables which contains sufficiently many A-variables. If the numbers of A- and Z-variables are relatively large, e.g., some $\frac{1}{c}$ fraction of the total number for a constant $C$, this appears to be quite plausible.

In this paper we were only able to give a similar type (having the addressing scheme) of Boolean functions for the separation. Seeking different types of Boolean functions having the same separation property will be also interesting.

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Abstract. We prove that a very basic class of program schemes augmented with access to a queue and an additional numeric universe within which counting is permitted accepts exactly the class of recursively solvable problems. The class of problems accepted when access to the numeric universe is removed is exactly the class of recursively solvable problems that are closed under extensions. We build upon NSPQ(1) an infinite hierarchy of classes of program schemes and show that the class of problems accepted by these program schemes has a zero-one law and consists of those problems defined in any vectorized Lindstr"om logic formed using operators whose corresponding problems are recursively solvable and closed under extensions. We apply our results to yield logical characterizations of complexity classes and provide logical analogs to inequalities and hypotheses from complexity theory involving the classes \text{NP} through to \text{ELEMENTARY}.

1 Introduction

Descriptive complexity theory essentially began with Fagin’s Theorem \cite{fagin74} which states that a problem can be accepted by a polynomial-time non-deterministic Turing machine iff it can be defined in existential second-order logic. Since then, the relative relationships of a whole range of complexity classes and logics have been examined (see, \textit{e.g.}, \cite{linial83, karp80}). Having logical equivalents of hypotheses in complexity theory enables attack using tools from finite model theory, such as Ehrenfeucht-Fra"issé games and zero-one laws.

The connection between a complexity class and a logic need not be as tight as in Fagin’s Theorem. For example, it is known that a problem is in \text{P} (resp. \text{PSPACE}) iff it can be defined in inductive (resp. partial) fixed-point logic, but in both cases only when we work with ordered finite structures (see \cite{linial83, karp80}). It turns out that on the class of all finite structures there are (computationally trivial) problems not definable in these fixed-point logics. Having logical characterizations of complexity classes on the class of all finite structures is preferable. For instance: ordered structures are more difficult to work with when applying the inexpressibility tools of finite model theory; in applications of descriptive complexity theory, \textit{e.g.}, to databases, it is almost always the case that finite structures are not ordered; and it is often undecidable as to whether a formula is well-formed when restricted just to ordered finite structures.
Whilst it remains unknown as to whether there is a logic capturing $P$ on the class of all finite structures, Abiteboul and Vianu proved the following result \cite{2,3}: $P = \text{PSPACE}$ iff inductive fixed-point logic and partial fixed-point logic are equally expressive on the class of all finite structures. It is particularly interesting that a complexity-theoretic statement like $P = \text{PSPACE}$ has an equivalent formulation in logic where there is a such a mismatch between the complexity class and the class of problems defined in the respective logic. Abiteboul, Vardi and Vianu \cite{1} and Dawar \cite{5} subsequently obtained similar results involving a range of complexity classes including $P$, $\text{NP}$, the Polynomial Hierarchy $\text{PH}$, $\text{PSPACE}$ and $\text{EXPTIME}$. One thing all of their logics have in common is that they can be realised as fragments of bounded-variable infinitary logic $L_{\omega_1\omega}$.

In obtaining their result, Abiteboul and Vianu developed a model of computation known as a loosely coupled generic machine (renamed a relational machine in \cite{1}). Abiteboul and Vianu are not the only researchers in recent times to work with models of computation taking finite structures as inputs. Perhaps the best known such computational model is the abstract state machine of Gurevich \cite{9}. Recently, the expressive power of different classes of program schemes has been considered. Program schemes are similar in flavour to relational machines and abstract state machines, and form a model of computation that is closer to the notion of a high-level program than a logical formula is; yet program schemes remain amenable to logical manipulation. In \cite{4,11,12,13}, the computational power of different classes of program schemes (on the class of all finite structures) is compared with the expressive power of certain logics. A crucial difference between program schemes and the traditional logics of descriptive complexity theory is that program schemes can be augmented with ‘computational’ constructs not really available in logic, such as stacks and arrays.

In this paper, we augment a very basic class of program schemes $\text{NPS}(1)$ with a queue and an extra ‘numeric’ universe. We prove that when both augmentations are present, the class of problems accepted by the resulting class of program schemes $\text{NPSQ}+_{\text{(1)}}$ is exactly the class of recursively solvable problems, \textit{i.e.}, the class of problems $\text{REC}$; and when only the queue is added, the class of problems accepted by the resulting class of program schemes $\text{NPSQ}(1)$ is exactly the class of recursively solvable problems that are closed under extensions, \textit{i.e.}, the class of problems $\text{REC} \cap \text{EXT}$. We proceed to introduce universal quantification into our program schemes by constructing an infinite hierarchy $\text{NPSQ}$ of classes of program schemes obtained by interleaving the application of universal quantification with the constructs of the program schemes of $\text{NPSQ}(1)$. It turns out that the class of problems accepted by the program schemes of $\text{NPSQ}$ has an alternative ‘semantic’ characterization as the class of problems defined by the sentences of any vectorized Lindström logic for which the problem corresponding to the operator is recursively solvable and closed under extensions. It follows as a corollary that this class of problems has a zero-one law. We remark that by \cite{11}, there are problems in $\text{NPSQ}(1) \cap \text{NP}$ that are not definable in $L_{\omega_1\omega}$.

The fact that $\text{NPSQ}+_{\text{(1)}}$ is computationally complete enables us to restrict the resources (time and space) used by these program schemes so that we can
logically capture many complexity classes ranging from \( \text{NP} \) through to \( \text{ELEMENTARY} \). By looking at the respective classes of program schemes where access to the numeric universe is removed, we move from the complexity class to its fragment of problems closed under extensions. In doing so, we show that many complexity-theoretic inequalities, e.g., \( \text{NEXPTIME} \neq 2\text{-NEXPTIME} \), and hypotheses, e.g., \( \text{NEXPTIME} \neq \text{EXPSPACE} \), are equivalent to these inequalities and hypotheses for the fragments, e.g., \( \text{NEXPTIME} \cap \mathcal{E}' \neq 2\text{-NEXPTIME} \cap \mathcal{E}' \) and \( \text{NEXPTIME} \cap \mathcal{E}' \neq \text{EXPSPACE} \cap \mathcal{E}' \); and that these fragments can be logically captured (in the sense of Gurevich [8]): it is not immediately obvious that a semantically restricted complexity class such as \( \text{NEXPTIME} \cap \mathcal{E}' \) can be captured syntactically by a logic. This is interesting for the same reason that the equivalence results of Abiteboul at al. are interesting, i.e., there exist computationally trivial problems not in \( \text{NEXPTIME} \cap \mathcal{E}' \) and \( \text{EXPSPACE} \cap \mathcal{E}' \). Whilst our equivalence results are similar to those of Abiteboul et al., they involve different logics: recall, that there are problems in \( \text{NPSQ}(1) \cap \text{NP} \) not definable in \( \mathcal{L}_{\infty\omega}^{\infty} \). Also, whereas the logics involved in the results due to Abiteboul el al. have a zero-one law (because they are fragments of \( \mathcal{L}_{\infty\omega}^{\infty} \)), our classes of program schemes define problems closed under extensions and so have a ‘one-law’ (apart from the program scheme which accepts no finite structures over some fixed signature). Not all proofs are included due to space limitations.

2 Program Schemes with Queues

Our signatures consist of a (non-empty) finite tuple of relation symbols and constant symbols. A problem is an isomorphism-closed class of finite structures over some fixed signature. If \( A \) and \( B \) are both finite \( \sigma \)-structures such that the domain of \( A \) is contained in the domain of \( B \), i.e., \( |A| \subseteq |B| \), and \( B \) restricted to \( |A| \) is \( A \) then \( B \) is an extension of \( A \) and we write \( A \subseteq B \). If \( \Omega \) is a problem, over the signature \( \sigma \), for which: \( A \) and \( B \) are finite \( \sigma \)-structures; \( A \subseteq B \); and \( A \in \Omega \), imply that \( B \in \Omega \) then \( \Omega \) is said to be closed under extensions. The class of problems that are closed under extensions is \( \mathcal{E}' \). Our structures are always finite.

Program schemes are more ‘computational’ means for defining classes of problems than are logical formulae. A program scheme \( \rho \in \text{NPS}(1) \) involves a finite set \( \{x_1, x_2, \ldots, x_k\} \) of variables and is over a signature \( \sigma \). It consists of a finite sequence of instructions where the first instruction is ‘\( \text{input}(x_1, x_2, \ldots, x_m) \)’, for some \( m \leq k \), and the variables of \( \{x_1, x_2, \ldots, x_m\} \) are known as the input-output variables, with the remaining variables being the free variables. The remaining instructions are one of the following: an assignment instruction of the form ‘\( x_i := y' \)’, where \( y \) is a variable or a constant symbol; a guess instruction of the form ‘\( \text{guess } x_i \)’; a while instruction of the form ‘\( \text{while } \varphi \text{ do } \alpha_1; \alpha_2; \ldots; \alpha_q \text{ od} \)’, where \( \varphi \) is a quantifier-free first-order formula over \( \sigma \) and where each of \( \alpha_1, \alpha_2, \ldots, \alpha_q \) is another instruction of one of the forms given here; or an accept instruction accept or a reject instruction reject. A program scheme \( \rho \in \text{NPS}(1) \)
over \( \sigma \) with \( k - m \) free variables takes expansions \( A' \) of \( \sigma \)-structures with \( k - m \) constants as input and computes on \( A' \) in the obvious way except that: execution of an instruction `guess \( x_i \)` non-deterministically assigns an element of \(|A'|\) to \( x_i \); initially, every input-output variable is non-deterministically assigned a value from \(|A'|\); and if a computation encounters an accept or a reject instruction then the computation halts. Note that the value of a free variable never changes throughout a computation: the free variables appear as if they were constant symbols. The structure \( A' \) is accepted by \( \rho \), and we write \( A'|=\rho \), iff there exists a computation of \( \rho \) such that an accept instruction is reached. We can easily build the usual `if` and `if-then-else` instructions using while instructions: henceforth, we shall assume that these instructions are at our disposal. Note that the class of structures accepted by a program scheme of NPS(1) is a problem.

Whereas a stack and arrays were incorporated into the program schemes of NPS(1) in [4,11,12,13], let us now augment these program schemes with access to a queue, i.e., a first-in, first-out store. In addition to the above instructions, we include the instructions `push \( x_i \)` and `\( x_i := \text{pop} \)` although in this latter instruction we insist that \( x_i \) must be an input-output variable, whereas in the former it can also be a free variable. The first instruction takes the current value of the variable \( x_i \) and pushes this value onto the end of the queue (the value of \( x_i \) does not change and the length of the queue increases by 1), and the second takes the current value from the head of the queue and sets \( x_i \) to have this value (with the length of the queue decreasing by 1). Initially, the queue is empty and if ever we attempt to pop from an empty queue then that particular computation is deemed not to be accepting. We denote the program schemes of NPS(1) augmented with a queue by NPSQ(1).

We can also augment our program schemes with a numeric universe. That is, we can assume that the variables of a program scheme are of one of two types: variables of the first type, the element type, take values from the universe of the input structure (as they have done so far); and variables of the second type, the numeric type, take values from the numeric universe, namely \( \{0, 1, \ldots, n - 1\} \), where the input structure has size \( n \) (we assume that the universe of the input structure and the numeric universe are disjoint). We insist that all variables of numeric type must be input-output variables and that they are initialised to 0. Additionally, there are two constant symbols 0 and \( \text{max} \) which are always interpreted as the numbers 0 and \( n - 1 \). The instructions available to the variables of numeric type, \( t_1, t_2, \ldots, t_q \), say, are assignments of the form `\( t_i := t_j \)`, `\( t_i := 0 \)`, `\( t_i := \text{max} \)` and `\( t_i := t_j + 1 \)` (if the variable \( t_j \) has value \( n - 1 \) then execution of this instruction causes that particular computation to reject the input); and `\( t_i = t_j \)`, `\( t_i = 0 \)`, `\( t_i = \text{max} \)` and their negations can appear as atoms and negated atoms in quantifier-free first-order formulae used as tests in while instructions (these formulae might be combinations of atoms involving variables of both element and numeric type). The class of program schemes NPSQ(1) augmented with a numeric universe is denoted NPSQ+\((1)\). Again, the classes of structures accepted by the program schemes of NPSQ+\((1)\) are problems.
Finite structures are abstract objects and as such generally have no canonical representation as a string over \( \{0, 1\} \), i.e., as a bit-string. Nevertheless, we can still arrange for finite structures to be input to Turing machines by encoding them as strings (throughout, our Turing machines are non-deterministic). We say that a problem \( \Omega \) over \( \sigma \) is accepted by a Turing machine \( M \) if \( M \) accepts exactly those bit-strings encoding structures in \( \Omega \); and so, in particular, as to whether an encoding \( e_\sigma(A) \) of any \( \sigma \)-structure \( A \) is accepted by \( M \) is independent of the particular linear order of the elements chosen when encoding \( A \). We can easily choose our encoding scheme so that the length of \( e_\sigma(A) \) is independent of the particular linear ordering chosen; and if \( |A| = n \) then we denote this length by \( e_\sigma(n) \). Define \( \mathcal{REC} \) as the class of problems accepted by some Turing machine. Note that not every Turing machine accepts a problem and that \( \mathcal{REC} \) is not the class of recursive languages but the class of recursive problems. However, any (resp. recursive) language can be realised as a problem over the signature \( \sigma \) consisting of the binary relation symbol \( L \) and the unary relation symbol \( B \) (resp. accepted by a Turing machine). A string over \( \{0, 1\} \) can be considered as a \( \sigma \)-structure \( A \) (in fact, as a class of \( \sigma \)-structures) whose binary relation \( L^A \) describes a linear order, i.e., the bit positions of the string, and whose unary relation \( B^A \) details which bits are 0 and which bits are 1.

Throughout, we identify a class of program schemes with the class of problems accepted by those program schemes (the same goes for sentences of logics and Turing machines).

**Theorem 1.** A problem is accepted by a program scheme of \( NPSQ^+_{=1} \) iff it can be accepted by a Turing machine; that is, \( NPSQ^+_{=1} = \mathcal{REC} \).

**Proof.** Suppose that the problem \( \Omega \) is accepted by some (non-deterministic) Turing machine \( M \). We shall construct a program scheme \( \rho \) of \( NPSQ^+_{=1} \) which simulates \( M \). Without loss of generality, we can assume that: the Turing machine \( M \) has a two-way infinite work-tape and a read-only input-tape; the work-tape alphabet is \( \{0, 1, b\} \); and the input-tape has a left-hand marker symbol \( l \) and a right-hand marker symbol \( r \) which delimit the input string.

Essentially, the queue of our program scheme \( \rho \) will hold a description of the tapes of the Turing machine. We start with the work-tape. Let \( u \) and \( v \) be distinct elements of an input structure (we can guess these elements as the first act of our program scheme). In our description of the work-tape, we encode the work-tape symbol 0 as the triple \((u, u, u)\), the work-tape symbol 1 as the triple \((v, v, v)\) and the blank symbol \( b \) as the triple \((u, u, v)\); and we use the triples \((v, u, u)\) and \((v, v, u)\) as delimiters. Let the work-tape at some instantaneous description (ID) of \( M \) be of the form \( w_1, w_2, \ldots, w_i, \ldots, w_m \), reading from left to right where \( w_1 \) (resp. \( w_m \)) is the first blank symbol to the left (resp. right) of the work-tape head beyond which the work-tape is entirely blank. Furthermore, suppose that the work-tape head is scanning the symbol \( w_i \). The queue will consist of: the triple encoding \( w_1 \); the triple encoding \( w_2 \); \ldots; the triple \((v, u, u)\) (to denote that the head is scanning the next symbol); the triple encoding \( w_i \); \ldots; the triple encoding \( w_m \); and the triple \((v, v, u)\) (to denote that we have reached the end of our description of the work-tape). We shall describe the state
of the Turing machine $M$ in any ID using a constant number of variables of our program scheme $\rho$. Also, should our program scheme $\rho$ know which move of the Turing machine $M$ to simulate and the effect upon the work-tape and the work-tape head, we could easily simulate this move by making appropriate manipulations of the queue.

Consequently, we are left with the problem of dealing with the input tape. This is more complicated as initially our program scheme is not given an encoding of the input structure $\mathcal{A}$, which has size $n$, say: it simply has access to the raw finite structure itself. However, we can build our own encoding of $\mathcal{A}$. Prior to any simulation of the Turing machine $M$, our program scheme guesses a linear ordering on $|\mathcal{A}|$ and stores this linear ordering on the queue. Note that because we have access to the (disjoint) domain $\{0, 1, \ldots, n - 1\}$, the capacity to count using the elements of this domain and the integers $0$ and $n - 1$ (through the constants $0$ and $\text{max}$), we can ensure that our guessed linear order contains every element of $|\mathcal{A}|$ exactly once.

Suppose that our guessed linear order is $u_0, u_1, \ldots, u_{n-1}$. This linear order defines a concrete encoding of our input structure and it is this encoding which we assume is presented to the Turing machine $M$. We use our linear order to write a description of the input-tape of $M$ on our queue.

It is now straightforward, given descriptions of the input-tape and the work-tape on the queue, to simulate the computation of $M$. The converse, that any problem in $\text{NPSQ}_+(1)$ is in $\text{REC}$, is trivial.

The question remains as to what sort of problems can be accepted by program schemes of $\text{NPSQ}(1)$.

**Theorem 2.** A problem is accepted by a program scheme of $\text{NPSQ}(1)$ iff it can be accepted by a Turing machine and is closed under extensions; that is, $\text{NPSQ}(1) = \text{REC} \cap \text{EXT}$. 

*Proof.* Let $\mathcal{A}$ and $\mathcal{B}$ be $\sigma$-structures such that $\mathcal{A} \subseteq \mathcal{B}$. If $\mathcal{A}$ is accepted by some program scheme $\rho$ of $\text{NPSQ}(1)$ then we can mirror an accepting computation in a computation of $\rho$ on $\mathcal{B}$ (essentially, because our tests in while instructions are quantifier-free). Hence, $\text{NPSQ}(1) \subseteq \text{REC} \cap \text{EXT}$. 

Conversely, suppose that $\Omega$ is a problem, over the signature $\sigma$, in $\text{REC} \cap \text{EXT}$. In particular, $\Omega$ is accepted by some Turing machine $M$. Construct a program scheme $\rho'$ of $\text{NPSQ}(1)$ as in the proof of Theorem 1 except that when the linear order is guessed, the checks made are that every element appears in the linear order at most once and that every constant appears in the linear order. Whenever $\rho'$ makes a guess, include code to ensure that the guess is always an element in the domain of the linear order.

Suppose that the $\sigma$-structure $\mathcal{A}$ is in $\Omega$; that is, every encoding of $\mathcal{A}$ is accepted by $M$. Then there is a computation of $\rho'$ on input $\mathcal{A}$ which guesses a linear order whose domain is the whole of $|\mathcal{A}|$; and as a result, this computation of $\rho'$ on input $\mathcal{A}$ simulates that of $M$ on the respective encoding of $\mathcal{A}$. Thus, $\mathcal{A}$ is accepted by $\rho'$. 

Suppose that the $\sigma$-structure $\mathcal{A}$ is accepted by $\rho'$. Then there is an accepting computation where this computation guesses a linear order whose elements come from the subset $B$ of $|\mathcal{A}|$. Let $\mathcal{B}$ be the restriction of $\mathcal{A}$ to $B$ (and so $\mathcal{A}$ is an extension of $\mathcal{B}$). The accepting computation of $\rho'$ on input $\mathcal{A}$ simulates an accepting computation of $M$ on input the respective encoding of $\mathcal{B}$ (because any guess results in an element of $B$). Hence, this encoding of $\mathcal{B}$ is accepted by $M$ and thus $\mathcal{B} \in \Omega$. But $\Omega$ is closed under extensions and so $\mathcal{A} \in \Omega$. The result follows. \hfill \Box

\section{Zero-One Laws}

We can extend our class of program schemes NPSQ(1) to a hierarchy of classes of program schemes essentially by interleaving applications of universal quantifiers with the basic constructs of the program schemes of NPSQ(1). In more detail, assume that we have defined a class of program schemes NPSQ(2$^{m-1}$), for some $m \geq 1$, and that any program scheme has associated with it: a set of input-output variables; a set of free variables; and a set of bound variables.

**Definition 1.** Let the program scheme $\rho \in \text{NPSQ}(2m-1)$ be over the signature $\sigma$, and suppose that $\rho$ has: input-output variables $x_1, x_2, \ldots, x_k$; free variables $x_{k+1}, x_{k+2}, \ldots, x_{k+s}$; and bound variables $x_{k+s+1}, x_{k+s+2}, \ldots, x_{k+s+t}$. Let $x_{i_1}, x_{i_2}, \ldots, x_{i_p}$ be free variables of $\rho$. Then $\forall x_{i_1} \forall x_{i_2} \ldots \forall x_{i_p} \rho$ is a program scheme of NPSQ(2$m$), which we denote by $\rho'$, with: no input-output variables; free variables those of $\{x_{k+1}, x_{k+2}, \ldots, x_{k+s}\} \setminus \{x_{i_1}, x_{i_2}, \ldots, x_{i_p}\}$; and the remaining variables of $\{x_1, x_2, \ldots, x_{k+s+t}\}$ as its bound variables.

A program scheme such as $\rho'$ takes expansions $\mathcal{A'}$ of $\sigma$-structures $\mathcal{A}$ by adjoining $s-p$ constants as input (one for each free variable), and $\rho'$ accepts such an expansion $\mathcal{A'}$ iff for every expansion $\mathcal{A''}$ of $\mathcal{A'}$ by $p$ additional constants (one for each variable $x_{i_j}$, for $j \in \{1, 2, \ldots, p\}$), $\mathcal{A''} \models \rho$ (the computation on such an expansion $\mathcal{A''}$ always starts with the queue empty). \hfill \Box

**Definition 2.** A program scheme $\rho' \in \text{NPSQ}(2m-1)$, for some $m \geq 2$, over the signature $\sigma$, is defined exactly as is a program scheme of NPSQ(1) except that the test in any while instruction is a program scheme $\rho \in \text{NPSQ}(2m-2)$. The bound variables of $\rho'$ consist of the bound variables of any test in any while instruction; all free variables in any test in any while instruction are input-output or free variables of $\rho'$; and there may be other free and input-output variables (appearing in $\rho'$ at the ‘top level’ but not in any test). Of course, any free variable never appears on the left-hand side of an assignment instruction, in a guess instruction or in a pop instruction in $\rho'$.

Suppose that a program scheme $\rho' \in \text{NPSQ}(2m-1)$ has $s$ free variables. Then it takes expansions $\mathcal{A'}$ of $\sigma$-structures $\mathcal{A}$ by adjoining $s$ constants as input and computes on $\mathcal{A'}$ in the obvious way; except that when some while instruction is encountered, the test, a program scheme $\rho \in \text{NPSQ}(2m-2)$, is evaluated according to the expansion of $\mathcal{A'}$ by the current values of any relevant input-output variables of $\rho'$. In order to evaluate this test, the queue associated with $\rho$
is initialised as empty and when the test has been evaluated the computation of $\rho'$ resumes accordingly with its queue and input-output and free variables being as they were immediately prior to evaluation.

So, we obtain a hierarchy $\text{NPSQ}(1) \subseteq \text{NPSQ}(2) \subseteq \ldots \subseteq \bigcup \{\text{NPSQ}(i) : i = 1, 2, \ldots \} = \text{NPSQ}$. Similar hierarchies have been defined previously in \cite{4,11,13} but based around stacks and arrays and not queues. We omit the proofs of the following two results but remark that the proof of the latter result follows from the proof of the former.

**Theorem 3.** The class of program schemes $\text{NPSQ}$ has a zero-one law.

Then the logic formed by extending first-order logic FO with a *vectorized sequence of Lindström quantifiers* corresponding to the problem $\Omega$ is denoted $(\pm \Omega)^*[\text{FO}]$: see \cite{6} for definitions. Such a logic is a *vectorized Lindström logic*.

**Corollary 1.** $\text{NPSQ} = \bigcup \{(\pm \Omega)^*[\text{FO}] : \Omega \in \text{REC} \cap \text{EXT}\}$.

Essentially, Corollary 1 equates the ‘syntactically-defined’ class of problems NPSQ with the ‘semantically-defined’ class of problems definable using first-order constructs in tandem with ‘recursive extension-closed’ operators. Whilst NPSQ is a natural sub-class of a computationally complete class of program schemes, there are problems in \text{REC} which have a zero-one law but which are not in NPSQ, one such being the ‘Hamiltonian cycle’ problem.

### 4 A Simple Application

It is a question of great importance in finite model theory and database theory as to whether there is a logic which captures exactly the class of polynomial-time solvable problems. Of course, one has to be precise about what one means by a logic and Gurevich \cite{8} has formulated a definition which has been widely adopted. A *logic* $L$ is given by a pair of functions $(\text{Sen}, \text{Sat})$ satisfying the following conditions. The function $\text{Sen}$ associates with every signature $\sigma$ a recursive set $\text{Sen}(\sigma)$ whose elements are called $L$-sentences over $\sigma$. The function $\text{Sat}$ associates with every signature a recursive relation $\text{Sat}_\sigma(\mathcal{A}, \varphi)$, where $\mathcal{A}$ is a $\sigma$-structure and $\varphi$ is a sentence of $L$. We say that $\mathcal{A}$ *satisfies* $\varphi$ if $\text{Sat}_\sigma(\mathcal{A}, \varphi)$ holds. Furthermore, we require that $\text{Sat}_\sigma(\mathcal{A}, \varphi)$ iff $\text{Sat}_\sigma(\mathcal{B}, \varphi)$ when $\mathcal{A}$ and $\mathcal{B}$ are isomorphic $\sigma$-structures. As yet no-one has exhibited a logic (in Gurevich’s sense) to capture any mainstream complexity class (for which the expectation is that it is properly) contained in $\text{NP}$. Logical characterizations of complexity classes on the class of ordered finite structures usually suffer from the difficulty that deciding whether a formula of the logic is ‘order invariant’ is undecidable. It is straightforward to verify that the classes of program schemes NPSQ (together with the classes therein) and NPSQ$^+(1)$ are logics in the sense of Gurevich; and so, for example, Theorem 2 implies that the class of recursively-solvable problems that are closed under extensions can be logically captured. However,
by imposing resource conditions upon our program schemes, we can logically capture other classes of problems.

A problem $\Omega$, over some signature $\sigma$, is accepted by a non-deterministic Turing machine $M$ in time $f(x)$ (resp. using space $g(x)$) if $M$ accepts $\Omega$ and for every encoding $e_\sigma(A)$ of any $\sigma$-structure $A$ of size $n$ in $\Omega$, there is an accepting computation of $M$ on input $e_\sigma(A)$ taking time at most $f(e_\sigma(n))$ (resp. using space at most $g(e_\sigma(n))$). In contrast, the time and space complexities of a program scheme of NPSQ$_+(1)$ or NPSQ(1) on some input structure are measured in terms of the size of the input structure. In order to evaluate the time taken or the space used in some computation of a program scheme of NPSQ$_+(1)$ or NPSQ(1) on some input structure of size $n$, we treat the execution of every assignment, quantifier-free test, pop, push, etc., as taking one unit of time, and we regard a variable or a place on the queue as occupying one unit of space. As for a non-deterministic Turing machine, the complexity of a program scheme of NPSQ$_+(1)$ or NPSQ(1) is derived using accepting computations only.

For any function $f(x)$, let $\mathcal{P}\mathcal{O}\mathcal{L}(f(x))$ denote $\{p_0(x).f(q_0(x)))^k + p_1(x). f(q_1(x))^k - 1 + \ldots + p_{k-1}(x)f(q_{k-1}(x)) + p_k(x) : k \geq 0, p_0, p_1, \ldots, p_k, q_0, q_1, \ldots, q_{k-1}$ are polynomials with integer coefficients$. If $C$ is a class of functions then $\mathcal{P}\mathcal{O}\mathcal{L}(C) = \bigcup_{f(x) \in C} \mathcal{P}\mathcal{O}\mathcal{L}(f(x))$. Denote the class of non-deterministic Turing machines that accept problems in time (resp. using space) $f(x)$ where $f(x)$ is a function from the class of functions $C$ as NTM$^f(C)$ (resp. NTM$^s(C)$), with NPSQ$_t^s(1)(C)$ and NPSQ$_s^s(1)(C)$ (resp. NPSQ$^s(1)(C)$ and NPSQ$_s^s(1)(C)$) defined likewise. The proofs of Theorems 1 and 2 yield the following.

**Corollary 2.** For any class of functions $C$:

\[- NTM^t(\mathcal{P}\mathcal{O}\mathcal{L}(C)) = NPSQ^t(1)(\mathcal{P}\mathcal{O}\mathcal{L}(C))
\[- NTM^t(\mathcal{P}\mathcal{O}\mathcal{L}(C)) \cap \mathcal{E}\mathcal{X}\mathcal{T} = NPSQ^t(1)(\mathcal{P}\mathcal{O}\mathcal{L}(C))
\[- NTM^s(\mathcal{P}\mathcal{O}\mathcal{L}(C)) = NPSQ^s(1)(\mathcal{P}\mathcal{O}\mathcal{L}(C))
\[- NTM^s(\mathcal{P}\mathcal{O}\mathcal{L}(C)) \cap \mathcal{E}\mathcal{X}\mathcal{T} = NPSQ^s(1)(\mathcal{P}\mathcal{O}\mathcal{L}(C)). \]

We can regard NPSQ$^t(1)(\mathcal{P}\mathcal{O}\mathcal{L}(C))$, for example, as a logic (in Gurevich’s sense) by identifying program schemes from this class with pairs $(\rho, F(x))$, where $\rho$ is a program scheme of NPSQ(1) and $F(x)$ is a function from $\mathcal{P}\mathcal{O}\mathcal{L}(C)$. Of course, this requires that the functions of $C$ form a recursive set and that there is a Turing machine which when given a function $f(x)$ from $C$ and a positive integer $n$, computes the value $f(n)$. Thus, Corollary 2 can be used to logically capture a whole host of complexity classes including NP, PSPACE, NEXPTIME, EXPSPACE, 2-NEXPTIME, 2-EXPSPACE, ..., ELEMENTARY. However, Corollary 2 can also be used to logically characterize ‘fragments’ of (the above) complexity classes obtained by intersecting the complexity class with the class of problems $\mathcal{E}\mathcal{X}\mathcal{T}$. Note that this is a ‘semantic’ rather than a ‘syntactic’ restriction and so it is not immediately obvious as to whether such a fragment of a complexity class can be logically captured. Furthermore, existing complexity-theoretic inequalities or hypotheses can be used to derive analogous ones for the respective fragments of the complexity classes in question.
Applying Corollary 2, we know that $\text{NEXPTIME} \neq 2\text{-NEXPTIME}$. Corollary 2 implies that $\text{NEXPTIME} \cap \mathcal{E} \mathcal{A} \mathcal{T} \neq 2\text{-NEXPTIME} \cap \mathcal{E} \mathcal{A} \mathcal{T}$ and that the class of problems accepted by program schemes of NPSQ(1) restricted to run in time $2^{2^{p(x)}}$, for some polynomial $p(x)$, properly contains the class of problems accepted by program schemes of NPSQ(1) restricted to run in time $2^{p(x)}$, for some polynomial $p(x)$. Also, it is open as to whether $\text{NEXPTIME} = \text{EXPSPACE}$, which, by Corollary 2, is equivalent to whether $\text{NEXPTIME} \cap \mathcal{E} \mathcal{A} \mathcal{T} = \text{EXPSPACE} \cap \mathcal{E} \mathcal{A} \mathcal{T}$ and to whether the class of problems accepted by program schemes of NPSQ(1) restricted to run in time $2^{p(x)}$, for some polynomial $p(x)$, is the same as the class of problems accepted by program schemes of NPSQ(1) restricted to run using space $2^{p(x)}$, for some polynomial $p(x)$.

We feel that it is interesting that open complexity-theoretic questions are equivalent to analogous logical questions involving classes of problems with zero-one laws, and even closed under extensions. Of course, these latter classes of problems do not contain some computationally-trivial problems, yet still they have a definitive role to play in the question of whether two complexity classes, such as $\text{NEXPTIME}$ and $\text{EXPSPACE}$ which have significant complexity-theoretic capacities, are identical.

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Algebraic Properties for P-Selectivity

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Abstract. Karp and Lipton, in their seminal 1980 paper, introduced the notion of advice (nonuniform) complexity, which since has been of central importance in complexity theory. Nonetheless, much remains unknown about the optimal advice complexity of classes having polynomial advice complexity.

In particular, let P-sel denote the class of all P-selective sets [23] For the nondeterministic advice complexity of P-sel, linear upper and lower bounds are known [10]. However, for the deterministic advice complexity of P-sel, the best known upper bound is quadratic [13], and the best known lower bound is the linear lower bound inherited from the nondeterministic case. This paper establishes an algebraic sufficient condition for P-sel to have a linear upper bound: If all P-selective sets are associatively P-selective, then the deterministic advice complexity of P-sel is linear. (The weakest previously known sufficient condition was P = NP.)

Relatedly, we prove that every associatively P-selective set is commutatively, associatively P-selective.

1 Introduction

Selman [23,24,25,26] defined the P-selective sets about twenty years ago. In addition to being of interest in their own right, they have recently had some surprising applications. For example, selectivity is a powerful tool in the study of search versus decision problems [4], and nondeterministic generalizations of selectivity are the key tools used to show that even NP machines cannot uniquely refine satisfying assignments unless the polynomial hierarchy collapses [8], that even weaker refinements are also precluded unless the polynomial hierarchy collapses [21,18], and that many cardinality types of nondeterministic function classes cannot collapse unless the polynomial hierarchy collapses [9].

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Definition 1 \cite{23}. A set $B$ is P-selective if and only if there is a (total) polynomial-time function $f : \Sigma^* \times \Sigma^* \rightarrow \Sigma^*$ such that

1. $(\forall x, y)[f(x, y) = x \lor f(x, y) = y]$. \\
2. $(\forall x, y)[(x \in B \lor y \in B) \implies f(x, y) \in B]$. \\

We call such an $f$ a P-selector function for $B$.

That is, a set $B$ is P-selective if there is a polynomial-time function $f$ that given any two strings always chooses one of them, and $f$ does this in such a way that if exactly one of the strings is in $B$ then the P-selector function chooses that string. $f(x, y)$ is often described, very informally, as choosing one of $x$ or $y$ that is “more likely (actually, no less likely) to be in the set.”

The P-selective sets have been extensively studied, and much about them is well understood. Though some P-selective sets are very complex—highly undecidable—the P-selective sets nonetheless have a broad range of structural simplicity properties. Most crucially in terms of the study in this paper, Ko \cite{13} showed that they have low nonuniform complexity ($\text{P-sel} \subseteq \text{P}/\text{O}(n^2)$). A few among the many other simplicity results known to hold are: Ko and Schöning \cite{15} showed that all P-selective sets in NP are in the second level of the low hierarchy of Schöning \cite{22}, and Allender and Hemachandra \cite{1} showed that the Ko-Schöning result is the strongest lowness result for P-sel that holds with respect to all oracles; a long line of work starting with Selman \cite{23} and Toda \cite{28} has shown that no P-selective set can be NP-hard under $\leq_p$ or various other reductions unless $\text{P} = \text{NP}$; Naik and Selman \cite{19} have shown that no P-selective set can be truth-table-hard for NP unless certain (intuitively unlikely) containments hold in the relationship between adaptive and nonadaptive queries to NP; and as a consequence of the work of Ko \cite{13} and Köbler and Watanabe \cite{16}, no P-selective set can be truth-table-hard for NP unless the polynomial hierarchy collapses to $ZPP^{NP}$, where ZPP as usual denotes expected polynomial time. Though much is known about the P-selective sets and their refinements and generalizations, including quite current work (e.g., \cite{20}), much remains open.

In this paper, we show that P-selector functions with nice algebraic properties have simplicity properties far beyond those known to hold for general P-selective sets. More generally, we study the class of languages one obtains from P-selector functions having algebraic properties, and the possibility of obtaining such algebraically nice P-selector functions.

In particular, Section 3 shows that any P-selective set having an associative P-selector function also has a commutative, associative P-selector function. It follows that sets having an associative P-selector can be accepted by deterministic advice interpreters using only a linear amount of advice. In contrast, the best upper bound on the deterministic advice complexity of the P-selective sets is the quadratic bound obtained by Ko \cite{13}.

Our result provides a new sufficient condition—all P-selective sets are associatively P-selective—for all P-selective sets having linear deterministic advice; the weakest previously known sufficient condition was the extremely demanding assumption that $\text{P} = \text{NP}$. 
Section 4 shows that associatively P-selective sets with weak census functions can never be coNP-immune. Section 5 establishes a structural sufficient condition for all P-selective sets being associatively, commutatively P-selective.

2 Definitions

A binary function $f$ is associative if and only if, for all $a$, $b$, and $c$, it holds that $f(a, f(b, c)) = f(f(a, b), c)$. A binary function $f$ is associative at each length if, for all $a$, $b$, and $c$ satisfying $|a| = |b| = |c|$, it holds that $f(a, f(b, c)) = f(f(a, b), c)$. A binary function $f$ is commutative if and only if, for all $a$ and $b$, $f(a, b) = f(b, a)$.

We will use the following notational shorthands for the classes that we will study. (The ordering of “A” and “C” changes to avoid confusion with the existing class “AC,” and to make clear that the “ℓ” modifies just the “A.”)

**Definition 2.**

1. $P$-sel = $\{B \mid B$ is P-selective$\}$.
2. $A$-sel = $\{B \mid B$ is P-selective via an associative P-selector function$\}$.
3. $C$-sel = $\{B \mid B$ is P-selective via a commutative P-selector function$\}$.
4. $CA$-sel = $\{B \mid B$ is P-selective via a commutative, associative P-selector function$\}$.
5. $A\ell$-sel = $\{B \mid B$ is P-selective via a P-selector function that is associative at each length$\}$.
6. $A\ell C$-sel = $\{B \mid B$ is P-selective via a commutative P-selector function that is associative at each length$\}$.

From these definitions, $CA$-sel ⊆ $A$-sel ⊆ $A\ell$-sel ⊆ $P$-sel, $CA$-sel ⊆ $A\ell C$-sel ⊆ $A\ell$-sel ⊆ $C$-sel ⊆ $P$-sel. The following fact is well-known.

**Fact 1** $P$-sel = $C$-sel.

The fact holds as if $B$ is P-selective via P-selector function $f$, then $f'(x, y) = \max(f(x, y), f(y, x))$ is a commutative function that remains a P-selector for $B$.

Advice classes capture the information content of sets with respect to some class of decoder (usually P).

**Definition 3 ([12]).**

1. Let $C$ be any complexity class and $f$ be any function mapping from $\mathbb{N}$ to $\mathbb{N}$. A set $B$ is in $C/f(n)$ if and only if there exist a set $C \in C$ and a function $h : 1^* \rightarrow \Sigma^*$ such that for all $x \in \Sigma^*$,
   (a) $|h(1^{|x|})| = f(|x|)$, and
   (b) $x \in B \iff \langle x, h(1^{|x|}) \rangle \in C$.

2. For any complexity class $C$ and any function class $F$, $C/F = \bigcup_{f \in F} C/f(n)$.

In this paper, we will be particularly interested in $P/O(n^2)$, $P/O(n)$, $P/n+1$, and $P/n$.

Let $F$ be any function complexity class. A set $B$ is $F$-printable if there is a function $g \in F$ such that, for all $x$, $g(x)$ outputs (in some fixed, standard
way of encoding sets as strings) \( B \cap (\Sigma^0 \cup \Sigma^1 \cup \cdots \cup \Sigma^{|x|}) \). That is, there is a function from \( F \) that finds all elements in the set up to a given length. By tradition, P-printable denotes the FP-printable sets.

As is standard, FP is the class of all (total) polynomial-time computable functions. FP\(^{\text{NP}}\) denotes the class of all functions that are computable in polynomial-time with the help of an NP oracle.

### 3 Associativity Drops Deterministic Advice Complexity to Linear

Are all associatively P-selective sets commutatively, associatively P-selective? If they are, this will not only collapse two of our classes, but also will allow us to use results from tournament theory in our study of associative selector functions (since, as we will discuss later, commutative selector functions have a close relationship to tournaments). Unfortunately, Fact 1 does not say that \( \text{A-sel} = \text{CA-sel} \). The reason it does not guarantee this is that it is conceptually possible that the transformation from \( f \) to \( f' \) that is stated just after Fact 1 will, while gaining commutativity, destroy associativity. The following theorem, our main result, states that this is never the case. (All proofs here omitted due to space can be found in [6].)

**Theorem 2.** \( \text{A-sel} = \text{CA-sel} \).

**Theorem 3.** \( \text{A}_\ell \text{-sel} = \text{A}_\ell \text{C-sel} \).

Ko showed the following result.

**Theorem 4 ([13]).** \( \text{P-sel} \subseteq \text{P/}O(n^2) \).

That is, deterministic advice interpreters given quadratic advice accept the P-selective sets. It is also known that nondeterministic advice interpreters with linear advice accept the P-selective sets, and this is optimal.

**Theorem 5 ([10]).** \( \text{P-sel} \subseteq \text{NP/n+1 but P-sel} \not\subseteq \text{NP/n} \).

It is natural to wonder whether those results can be unified into the very strong claim: \( \text{P-sel} \subseteq \text{P/}O(n) \). Currently, no proof of this claim is known and people doing research on selectivity generally doubt that it holds. However, proving the claim false seems unlikely in the immediate future, as by Theorem 5 any such proof would implicitly prove \( \text{P} \not= \text{NP} \).

Nonetheless, the following result shows that all associatively P-selective sets are accepted by deterministic advice interpreters with linear advice.

**Theorem 6.** \( \text{A}_\ell \text{-sel} \not\subseteq \text{P/n+1} \).

**Corollary 1.** \( \text{A-sel} \not\subseteq \text{P/n+1} \).
Proof of Theorem 6: Let $B \in A_\ell$-sel. Let $f$ be a commutative P-selector for $B$ that is associative at each length. By Theorem 3 we know that such a commutative P-selector for $B$ must exist.

Let $n \in \mathbb{N}$. Consider the directed graph $G_n = (B^{-n}, E_n)$, where $E_n = \{(x, y) \mid x, y \in B^{-n} \land x \neq y \land f(x, y) = y\}$. $G_n$ is a tournament—a directed graph such that between any two nodes $u, v, u \neq v$, there is exactly one of the edges $(u, v)$ or $(v, u)$.

$E_n$ is a transitive relation. To see this, let $x, y, z \in B^{-n}$. Suppose $(x, y) \in E_n$ and $(y, z) \in E_n$. Note that this implies $x \neq y$, $y \neq z$, and $x \neq z$. It also implies that $f(x, y) = y$ and $f(y, z) = z$. Thus, $f(f(x, y), z) = z$ and $f(x, f(y, z)) = f(x, z)$. Since $f$ is associative at each length we have $f(f(x, y), z) = f(x, f(y, z))$ and so $f(x, z) = z$, which implies $(x, z) \in E_n$. Thus, $E_n$ is transitive.

Tournaments such that their edge set is a transitive relation are called transitive tournaments. We will use the following result (which can be viewed as a tournament-language expression of the fact that each finite linear order has a minimal element).

Proposition 1 ([17]). The following statements are equivalent for tournaments $G = (V, E)$:

1. $G$ is a transitive tournament.
2. $G$ contains no directed cycles.
3. Every induced (directed) subgraph $G' = (V', E')$, $V' \neq \emptyset$, of $G$ contains a (distance-one) source node, i.e., a node $u \in V'$ such that for all $v \in V'$, $v \neq u$, $(u, v) \in E'$.

Since $f$ is associative at each length, we have for all $n \in \mathbb{N}$ that $G_n$ is a transitive tournament, which in turn by Proposition 1 implies that (if $G_n$ is not the empty tournament) $G_n$ contains a source node $s_n$. In other words, for all $n \in \mathbb{N}$, if $B^{-n} \neq \emptyset$ then there exists a string $s_n \in B^{-n}$ such that for all $x \in B^{-n}$, $f(x, s_n) = x$. Since $s_n \in B^{-n}$, clearly for each length $n$ string $z \in B$ we have $f(z, s_n) = s_n$.

Define, for all $n \in \mathbb{N}$, $h(1^n) = 1s_n$ if $B^{-n} \neq \emptyset$, and $h(1^n) = 0^{n+1}$ otherwise. Note that $h$ on any input $1^n$ outputs a string of length exactly $n + 1$. Let $L = \{\langle x, 1y \rangle \mid |x| = |y| \land f(x, y) = x\}$. Clearly, $L \in \mathsf{P}$. Due to the above construction we also have, for all $x \in \Sigma^*$,

$$x \in B \iff \langle x, h(|x|) \rangle \in L.$$  

This shows that $B \in \mathsf{P}/n+1$.

It remains to show that $\mathsf{P}/n+1 - A_\ell$-sel $\neq \emptyset$. In fact, it clearly holds even that $\mathsf{P}/1 - \mathsf{P}$-sel $\neq \emptyset$. The proof of this claim consists of a diagonalization against all P-selectors. Let $f_1, f_2, \ldots$ be an enumeration (not necessarily effective, though in fact there do exist such effective enumerations) of all polynomial-time selector functions (i.e., all polynomial-time functions such that for all $x$ and $y$ it holds that $f(x, y) \in \{x, y\}$). Set $B = \{0^{2i-1} \mid f_i(0^{2i-2}, 0^{2i-1}) = 0^{2i-2}\} \cup \{0^{2i-2} \mid f_i(0^{2i-2}, 0^{2i-1}) = 0^{2i-1}\}$. $B$ is not P-selective, yet $B \in \mathsf{P}/1$. 

\[ \square \]
Regarding the previous result, note that, more generally, if a set \( A \) and a (not necessarily commutative) P-selector \( f \) for it have the property that at each nonempty length \( m \) there is some string \( y \) in the set at that length—or even at some length linearly related to that length—such that \( f(y, z) = z \) or \( f(z, y) = z \) for each \( z \) of length \( m \), then \( A \) is in \( \text{P}/\mathcal{O}(n) \). However, among ways of ensuring that this condition is met, we feel that associativity is a particularly natural and well-structured property to study.

The number of advice bits in Theorem 6 and Corollary 1, \( n + 1 \), is optimal.

Theorem 7. \( \text{A-sel} \not\subseteq \text{P}/n \).

Corollary 2. \( \text{A}_\ell\text{-sel} \not\subseteq \text{P}/n \).

In fact, for any recursive function \( f \), \( \text{A-sel} \not\subseteq \text{DTIME}[f(n)]/n \). The reason for this is that we may take the P-selective set that Hemaspaandra and Torenvliet [10] prove is not in \( \text{DTIME}[f(n)]/n \) and, just as in the proof of Theorem 7, may complete its P-selector function in a manner that achieves associativity.

We mention in passing that—since \( \log_2((2^n+1-1)+1) = n+1 \) and thus one can include in the advice-seeking tournament all lengths up to the current one—the \( n + 1 \) advice bounds in Theorem 6 and Corollary 1 hold even in the “strong advice” model of Ko, Balcázar, and Schöning [14,2] (see also the discussion in [10]).

We note that our results relativize. In particular, one can relativize them by any particular \( \text{NP} \cap \text{coNP} \) set, and so regarding associativity and the \( \text{NP} \)-selective sets (equivalently the \( \text{NPSV}_1 \)-selective sets and equivalently the \( \text{FP}_{\text{NP} \cap \text{coNP}} \)-selective sets, see [8,7] for definitions and discussion), it follows from our results that all length-associatively \( \text{NP} \)-selective sets are in \( (\text{NP} \cap \text{coNP})/n + 1 \). In fact, essentially all of this paper’s results can be stated and proven more generally so as to apply even to a variety of nondeterministic selectivity classes; see [5] for full details.

4 Printability and Nonimmunity

Associativity yields additional simplicity properties. Let us consider nonimmunity results (i.e., presence of infinite subsets). For general \( \text{P} \)-selective sets, Theorem 8 holds. In contrast, for associatively selective sets, we have Theorem 9. To see how these bounds relate, note that \( \text{UP}_{\text{NP}} \subseteq \text{NP} \).

Theorem 8. Every infinite \( \text{P} \)-selective set \( B \) has an infinite \( \text{FP}^{\text{B} \oplus \text{NP}^{\text{NP}}} \)-printable subset.

Theorem 9. Every infinite \( \text{A} \)-selective (or even \( \text{A}_\ell \)-selective) set \( B \) has an infinite \( \text{FP}^{\text{B} \oplus \text{UP}^{\text{NP}}} \)-printable subset.

Theorem 8 should be contrasted with the result of Hemaspaandra, Ogihara, Zaki, and Zimand (personal communication, 1999) that \( \text{P-sel} \) is not (weak-\( \text{FP}^{\text{NP}^{\text{NP}}} \)-rankable)-immune, a result related to the study in [11].
Theorem 11. Every
then these classes coincide.
whether in fact A-sel is all of P -sel. We will prove, as Corollary 3, that if P = NP,
Many of our results give simplicity properties of A-sel. It is natural to wonder
5 Are All P-Selective Sets Associatively P-Selective?
Proof of Theorem 9: Let B be an infinite Aℓ-selective set. Let f be a commutative P-selector of B that is associative at each length. By Theorem 3 we know that such a selector for B exists.
Consider the function score that is defined as score(x) = ||{z ∈ Σ|x| | f(x, z) = x}||. Note that for each n ∈ N and for each x ∈ Σn, (a) x /∈ B → score(x) ≤ 2n − ||B−n|| and (b) x ∈ B → score(x) ≥ 2n − ||B−n|| + 1. So for all x ∈ B and all y /∈ B such that |x| = |y|, score(y) < score(x). It follows from the associativity-at-each-length of f that for all n, max{score(y) | |y| = n} = 2n (see also Proposition 11 and the comment immediately after it, applying both to the
tournament on Σn defined by the associativity-at-each-length selector function f).
So for each n, there exists exactly one string at length n, call it d_n, at length n, such that score(d_n) = max{score(y) | |y| = n} = 2n. Thus, the set
F = \{⟨1^n, z⟩ | n ∈ N ∧ z ∈ Σ* ∧ (∃ω : |ω| = n − |z|)[score(zω) = 2^n]\}
is in UPNP. So, given 1^n, computing d_n can be done in polynomial time with the help of the oracle F. Let C = B ∩ {d_i | i ≥ 0}. Clearly, as B is infinite, and C is both infinite and FP^{B\oplus UPNP}-printable.

If one cares about the number of queries to B needed in the printability claims of Theorems 8 and 9 the following can be observed. As communicated to us by Till Tantau [27], one can limit one’s queries to B to a logarithmic number via putting the queries into a Toda-like ordering and binary searching to find which are in and which are out.

It follows from the proof of Theorem 9 that any associatively selective set for which there is an infinite set of lengths at which we know it is not empty is in fact non-coNP-immune. This is a very weak type of partial census information.

Definition 4. We say a set B is hintable if there is a tally set, T ⊆ 1*, such that T ∈ P and (∀i)[1^i ∈ T ⟹ ||B^{=i}|| ≠ 0].

Theorem 10. Any infinite, hintable A-selective (or even Aℓ-selective) set has an infinite coNP subset (i.e., Aℓ-sel ∩ Hintable is not coNP-immune).

5 Are All P-Selective Sets Associatively P-Selective?
Many of our results give simplicity properties of A-sel. It is natural to wonder whether in fact A-sel is all of P-sel. We will prove, as Corollary 3 that if P = NP, then these classes coincide.

Theorem 11. Every P-selective set has a selector function in FP^{NP} that is commutative and associative.

Proof: Let A be a P-selective set. Let f be a P-selector of A, without loss of generality assume f to be commutative. For every pair of strings x and y, ω is called a connector of x and y if and only if either (a) f(x, ω) = ω and f(ω, y) = y or (b) f(x, ω) = x and f(ω, y) = ω. Note that for any two strings x and y, there
always exist connectors of \( x \) and \( y \), namely \( x \) and \( y \). Define the function \( g \) by setting, for all \( x \) and \( y \),
\[
g(x,y) = \begin{cases} 
y & \text{if the lexicographically smallest connector } \omega \text{ of } x \text{ and } y \\
x & \text{otherwise.}
\end{cases}
\]

It is not hard to see that \( g \) is an \( \text{FP}^{\text{NP}} \) function. We will now argue that \( g \) is a commutative and associative selector for \( A \). By definition, \( g \) is clearly commutative.

We will show that \( g \) is a selector for \( A \) by showing that for any two strings \( x \) and \( y \), \( g(x,y) \) satisfies the conditions (1) \( g(x,y) = x \lor g(x,y) = y \) and (2) \( (x \in A \lor y \in A) \implies g(x,y) \in A \) (see Definition 1). Let \( x \) and \( y \) be any two strings. Condition (1) holds by definition of \( g \). If \( x = y \) or \( f(x,y) = g(x,y) \) then also (2) is trivially satisfied. So suppose \( x \neq y \) and \( f(x,y) \neq g(x,y) \). Without loss of generality assume \( f(x,y) = x \) and \( g(x,y) = y \). Hence the (lexicographically) smallest connector of \( x \) and \( y \), call it \( \omega \), yields \( f(x,\omega) = \omega \) and \( f(\omega,y) = y \).

Since \( f(x,y) = x \), \( f(x,\omega) = \omega \), and \( f(\omega,y) = y \) it follows from the fact that \( f \) is a selector of \( A \) that either \( x,y,\omega \in A \) or \( x,y,\omega \notin A \). Hence (2), i.e., \( (x \in A \lor y \in A) \implies g(x,y) \in A \), holds.

It remains to show that \( g \) is an associative function. Assume that \( g \) is not associative. Hence there exist three disjoint strings \( x \), \( y \), and \( z \) such that \( g(x,g(y,z)) \neq g(x,y,z) \). (Note that if \( x = y \) or \( y = z \) or \( x = z \) then \( g(x,g(y,z)) = g(x,y,z) \).) It is not hard to see that \( g(x,g(y,z)) \neq g(x,y,z) \) is equivalent to
\[
(g(x,y) = y \land g(y,z) = z \land g(x,z) = x) \lor (g(x,y) = x \land g(y,z) = y \land g(x,z) = z).
\]

Informally put, \( g(x,g(y,z)) \neq g(x,y,z) \) is equivalent to the statement that \( x \), \( y \), and \( z \) induce a directed cycle of length three in the (infinite) tournament induced by \( g \). Let \( \omega \) be the lexicographically smallest connector among all connectors of \( x \) and \( y \), among all connectors of \( y \) and \( z \), and among all connectors of \( x \) and \( z \). Let \( a,b \in \{x,y,z\} \), \( a \neq b \), be such that \( \omega \) is a smallest connector of \( a \) and \( b \). Without loss of generality assume \( g(a,b) = b \). Hence \( f(a,\omega) = \omega \) and \( f(\omega,b) = b \). Let \( c \) be the string in \( \{x,y,z\} \setminus \{a,b\} \). If \( f(c,\omega) = c \) then \( \omega \) is also the (lexicographically) smallest connector of \( a \) and \( c \) and thus \( g(a,c) = c \), contradicting (\(*\)). If \( f(c,\omega) = \omega \) then \( \omega \) is also the (lexicographically) smallest connector of \( b \) and \( c \) and thus \( g(b,c) = b \), contradicting (\(*\)). This shows that \( g \) is associative and completes our proof.

**Corollary 3.** \( \text{P = NP} \implies \text{P-sel = A-sel} \).

Theorem 3 provides a sufficient condition, based on an algebraic property for selector functions, for \( \text{P-sel} \subseteq \text{P/O}(n) \). If one were interested only in *structural complexity-class-collapse* sufficient conditions, Theorem 3 would be no improvement over the \( \text{P = NP} \) sufficient condition implicit in the result of Hemaspaandra.
and Torenvliet that P-sel ⊆ NP/O(n) (and thus P-sel ⊆ P/O(n) if P = NP), as in light of Corollary 3, the best known structural complexity-class-collapse condition sufficient to imply P-sel = A-sel is also the collapse P = NP. However, we feel that this is the wrong view, and that “P-sel = A-sel” is probably a fundamentally different type of assumption than P = NP. For example, if P-sel = A-sel were in fact equivalent to P = NP, then P = NP would (trivially) be not just a sufficient condition for P-sel = A-sel but also would be a necessary condition. In fact, not only is P = NP not known to be necessary for P-sel = A-sel, but in fact no structural complexity-class-collapse condition—not even very weak collapses like P = UP or P = ZPP—is known to be necessary. Our point here is that, though by Corollary 3 P = NP is one way to achieve P-sel = A-sel, we conjecture that it is unlikely to characterize P-sel = A-sel. And thus, our P-sel = A-sel sufficient condition for P-sel ⊆ P/O(n) is best viewed as a new algebraic sufficient condition quite different from the known (and extremely demanding) structural complexity-class-collapse sufficient conditions.

Finally, we mention two open issues. First, can one prove a complete or partial converse of Corollary 3? The second issue is the following. One would ultimately like to know whether all P-selective sets have linear deterministic advice, i.e., whether P-sel ⊆ P/O(n). This paper gave a new sufficient condition for that, namely, P-sel ⊆ P/O(n) if all P-selective sets have associative (or even merely length-associative) selector functions. On the other hand, the task of constructing an oracle relative to which P-sel ⊆ P/O(n) fails has resisted any progress for half a decade. This paper suggests a weaker question that can be tackled first: trying to construct an oracle relative to which some P-selective sets have no length-associative selector functions.

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References


Abstract. $P$-complete problems seem to have no parallel algorithm which runs in polylogarithmic time using a polynomial number of processors. A $P$-complete problem is in class $EP$ (Efficient and Polynomially fast) if and only if there exists a cost optimal algorithm to solve it in $T(n) = O(t(n)^\epsilon)$ ($\epsilon < 1$) using $P(n)$ processors such that $T(n) \times P(n) = O(t(n))$, where $t(n)$ is the time complexity of the fastest sequential algorithm which solves the problem. The goal of our research is to find $EP$ parallel algorithms for $P$-complete problems. In this paper we consider two $P$-complete geometric problems in the plane. First we consider the convex layers problem of a set $S$ of $n$ points. Let $k$ be the number of the convex layers of $S$. When $1 \leq k \leq n^{1-\epsilon}$ ($0 < \epsilon < 1$) we can find the convex layers of $S$ in $O(n \log n \frac{\log n}{p})$ time using $p$ processors, where $1 \leq p \leq n^{\frac{1-\epsilon}{2}}$. Next, we consider the envelope layers problem of a set $S$ of $n$ line segments. Let $k$ be the number of the envelope layers of $S$. When $1 \leq k \leq n^{\frac{1}{2}}$ ($0 < \epsilon < 1$), we propose an algorithm for computing the envelope layers of $S$ in $O(n^{\alpha(n)} \log^3 n \frac{\log n}{p})$ time using $p$ processors, where $1 \leq p \leq n^{\frac{1-\epsilon}{2}}$, and $\alpha(n)$ is the functional inverse of Ackermann’s function which grows extremely slowly. The computational model we use in this paper is the $EREW$-$PRAM$. Our first algorithm, for the convex layers problem, belongs to $EP$, and the second one, for the envelope layers problem, belongs to the class $EP$ if a small factor of $\log n$ is ignored.
processors. The class of $P$-complete problems consists of the most likely candidates of $P$ that are not in $NC$. However, polylogarithmic time complexity is not so important when considering practical parallel computation. Actually, the number of processors available is usually small in comparison with the size of a problem. Thus, in practice cost optimality turns to be the most important measure for parallel algorithms. The cost of a parallel algorithm is defined as the product of the running time and the number of processors required by the algorithm. A parallel algorithm is called cost optimal if its cost is of the same order as the time complexity of the fastest known sequential algorithm for the same problem. Kruskal et al. proposed class $EP$ (Efficient and Polynomially fast). A problem is in $EP$ if and only if there exists a cost optimal algorithm to solve it in $T(n) = O(t(n)^\epsilon)$ ($\epsilon < 1$) using $P(n)$ processors such that $T(n) \times P(n) = O(t(n))$, where $t(n)$ is the time complexity of the fastest sequential algorithm which solves the problem.

In this paper, we consider $EP$ algorithms for two $P$-complete geometric problems, the convex layers problem and the envelope layers problem on $EREW$-$PRAM$. Given a set $S$ of $n$ points in the plane, $S$ can be divided into a sequence $CL(S)$ of convex hulls: $CL_1(S), CL_2(S), \ldots, CL_k(S)$, where $CL_i(S)$ ($1 \leq i \leq k$) is the convex hull of $\bigcup_{j=i+1}^{k} CL_j(S)$. The envelope layers of a set $S$ of (opaque) $n$ line segments are analogous to the convex layers of a set of points, with convex hulls replaced by upper envelopes. The upper envelope of $S$ is the collection of segment portions visible from the point $(0, +\infty)$, which consists of $O(n\alpha(n))$ subsegments of $S$. These two problems are $P$-complete. They are still $P$-complete even if $1 \leq k \leq n^{\frac{1}{2}}$.

Chazelle proposed an optimal sequential algorithm for the convex layers problem which runs in $O(n \log n)$ time. In Fujiwara et al. considered the convex layers problem under a very strong constraint, in which they proved that if all points of $S$ are on $d$ horizontal lines, when $d \leq n^{\epsilon}$ ($0 < \epsilon \leq \frac{1}{2}$) the problem is still $P$-complete. They proposed an $EP$ algorithm for the problem which runs in $O\left(\frac{n \log n}{p}\right)$ time using $p$ processors if $d \leq n^{\epsilon}$ ($0 < \epsilon \leq \frac{1}{2}$) and $1 \leq p \leq n^{\epsilon}$ in the $EREW^p$-$PRAM$. That is, to achieve cost optimality, there must be $d$ horizontal lines such that each line must pass through more than $n^{\frac{1}{2}}$ points in average, what is unlikely in most cases. Besides, the parameter $d$ does not represent the substantial complexity of the problem. In this paper we present a new $EP$ parallel algorithm for computing the convex layers of a set $S$ of $n$ points. Let $k$ be the number of the convex layers of $S$. When $1 \leq k \leq n^{\frac{1}{2}}$ ($0 < \epsilon < 1$) our algorithm runs in $O\left(\frac{n \log n}{p}\right)$ time using $p$ processors, where $1 \leq p \leq n^{\frac{1}{2} - \epsilon}$, in the $EREW^p$-$PRAM$, and it is cost optimal.

Let $S$ be a set of $n$ (opaque) line segments in the plane. Hersberger gave an $O(n\alpha(n) \log^2 n)$ algorithm for computing the envelope layers of $S$, where $\alpha(n)$ is the functional inverse of Ackermann’s function which grows extremely slowly. Here, we also give an algorithm for the envelope layers problem. Let $k$ be the number of envelope layers of $S$. When $0 \leq k \leq n^{\frac{1}{2}}$ ($0 < \epsilon < 1$) our algorithm runs in $O\left(\frac{n\alpha(n) \log^3 n}{p}\right)$ time using $p$ processors, where $1 \leq p \leq n^{\frac{1}{4} - \epsilon}$, in the $EREW^p$-$PRAM$. If we ignore a factor of $\log n$ our algorithm belongs to the class
In the following sections, we give the outline of our algorithms. The details can be found in [2].

2 Finding Convex Layers of Points

In our algorithm the following operations are required: (I) storing a convex layer, (II) searching for an edge of a convex layer, (III) finding the common tangents of two convex layers, (IV) deleting a set of contiguous edges from a convex layer, (V) given two convex polygons \( P \) and \( Q \) finding the convex hull of \( P \) and \( Q \), and (VI) given \( x \) convex polygons, \( P_1, P_2, \ldots, P_x \) with \( f \) vertices each, finding the common tangents of \( P_i \) and \( P_j \) for all pairs of \( i, j \) (\( 1 \leq i < j \leq x \)). Let the sizes of the convex layers and convex polygons are \( O(n) \). We keep convex layers in balanced tree. Operation (I) can be done in \( O(\log n) \) using \( n \) processors [7], operation (II) and (III) can be done in \( O(\log n) \) time sequentially [9]. Operations (IV) and (V) can be done in \( O(\log n) \) time sequentially [2]. Operation (VI) can be done in \( O(\frac{\log x \log^2 f x^2}{\log^2 x}) \) \( (p \leq fx^2) \) time using \( p \) processors [8]. Let \( S \) be a set of \( n \) points in the plane and \( k \) be the number of the convex layers of \( S \). We sort \( S \) by its \( x \)-coordinates. The following algorithm finds the convex layers of \( S \).

**Algorithm ComputeCL(S)**

- **[Input]** A set \( S = (p_1, p_2, \ldots, p_n) \) of \( n \) points in the Euclidean plane sorted by their \( x \)-coordinates.
- **[Output]** A set \( CL(S) = (CL_1(S), CL_2(S), \ldots, CL_k(S)) \) \( (1 \leq k \leq n) \) of the convex layers of \( S \), where \( CL_i(S) \) \( (1 \leq i \leq k) \) is the \( i \)th convex layer of \( S \).

**Phase 1** Divide \( S \) into \( S_1, S_2, \ldots, S_{\frac{1}{n+x}} \) subsets such that \( S_i \) \( (1 \leq i \leq \frac{1}{n+x}) \) contains \( n \frac{1}{x} \) points and the \( x \)-coordinate of any point of \( S_i \) is less than the \( x \)-coordinate of any point of \( S_{i+1} \).

**Phase 2** In parallel, for each subset \( S_i \) \( (1 \leq i \leq \frac{1}{n+x}) \) compute the convex layers of \( S_i \), denoted as \( CL(S_i) = (CL_1(S_i), CL_2(S_i), \ldots, CL_{k_i}(S_i)) \) by Chazelle’s sequential algorithm, where \( k_i \) is the number of layers in \( S_i \). Store each \( CL_j(S_i) \) \( (1 \leq j \leq k_i) \) \( (1 \leq i \leq \frac{1}{n+x}) \) in balanced tree.

**Phase 3** Let \( S'_i = S_i \) \( (1 \leq i \leq \frac{1}{n+x}) \). Let \( x = 1 \). Find \( CL_x(S) \), that is, the \( x \)th convex layer of \( S \), repeatedly as follows.

(a) Find the convex hull of the outermost layers of \( S_1', S_2', \ldots, S'_{\frac{1}{n+x}} \), which obviously is \( CL_x(S) \).

(b) Revise \( S'_i \) \( (1 \leq i \leq \frac{1}{n+x}) \) as follows: for each \( i \) \( (1 \leq i \leq \frac{1}{n+x}) \) in parallel, delete the vertices of \( CL_x(S) \) from \( S'_i \), and reconstruct the convex layers of \( S'_i \).

(c) Set \( x = x + 1 \). While all of \( S'_1, S'_2, \ldots, S'_{\frac{1}{n+x}} \) are not completely empty return to (a).

In the above algorithm, Phase 1 can be executed in \( O(1) \) time using \( n \) processors, if \( S \) is stored in an array. In Phase 2, \( CL(S_i) \) can be computed by Chazelle’s sequential algorithm [1] in \( O(n \frac{1}{x} \log n) \) time. Therefore, all the convex layers of \( S_1, S_2, \ldots, S'_{\frac{1}{n+x}} \) can be computed in \( O(n \frac{1}{x} \log n) \) time using \( n \frac{1}{x} \) processors.
We can store \( CL_j(S_i) \) \((1 \leq j \leq k_i, 1 \leq i \leq n^{1-\epsilon})\) in balanced tree in \(O(\log |S_i|)\) time using \(|S_i|\) processors. Now let us consider Phase 3. Phase 3(a) can be done in \(O(n)\) time using \(O(n^{1-\epsilon})\) processors[3]. Using the operations we defined, Phase 3(b) can be done in \(O(k \log n)\) time using \(O(n^{\frac{1}{k+\epsilon}})\) processors. Since the instructions of Phase 3 are repeated \(k\) times, where \(k\) is the number of convex layers of \(S\), Phase 3 takes totally \(O(k^2 \log n)\) time using \(O(n^{\frac{1}{k+\epsilon}})\) processors. Thus, the whole algorithm \(\text{ComputeCL}(S)\) can be executed in \(O(\max(n^{\frac{1}{k+\epsilon}} \log n, k^2 \log n))\) time using \(n^{\frac{1}{k+\epsilon}}\) processors. Thus by using \(p\) processors, where \(1 \leq p \leq n^{\frac{1}{k+\epsilon}}\), it takes \(O(\max(n \frac{\log n}{p}, n^{\frac{1}{k+\epsilon}} \frac{k^2 \log n}{p}))\) time in the \(EREW-PRAM\). Therefore, when \(1 \leq k \leq n^{\frac{1}{\epsilon}}\) \((0 < \epsilon < 1)\), the algorithm \(\text{ComputeCL}(S)\) is cost optimal and runs in \(O(\frac{n \log n}{p})\) time using \(p\) processors, where \(1 \leq p \leq n^{\frac{1}{k+\epsilon}}\).

3 Finding Envelope Layers of Line Segments

Let \(S\) be a set of \(n\) (opaque) line segments, and \(k\) be the number of the envelope layers of \(S\). From the duality of lines and points, the convex layers of lines can be found by the algorithm for the convex layers of points. The main idea of our algorithm here is to reduce the problem of line segments into that of lines by using a segment tree.

The plane can be partitioned into \(2n + 1\) slabs by drawing vertical lines through all the segments endpoints of \(S\). The segment tree of \(S\), denoted as \(ST(S)\), is built as follows[10]. Construct a complete binary tree with \(2n\) leaves. Then let each leaf of \(ST(S)\) represent one slab taken in left-to-right order, and each internal node represent the union of its descendants’ slabs. Whether original slabs or a union of them, are called canonical slabs. Descending from the root to the leaves, associate with each node \(v \in ST(S)\) a set \(S[v]\), where \(S[v]\) consists of the segments or subsegments of \(S\) that have its endpoints on the boundary of the canonical slab represented by node \(v\), which have not been associated with any of \(v\)’s ancestors in \(ST(S)\). For any node \(v\), the line segments in \(S[v]\) have the same left and right endpoints, respectively, therefore, they can be considered as lines. On the other hand, for any two nodes \(x\) and \(y\) on the same level of \(ST(S)\), the subsets associated with nodes \(x\) and \(y\) are separated. Thus, for the sets associated on the same level their envelopes can be found in parallel by the algorithm for the convex layers of points.

**Algorithm ComputeEL(S)**

[Input] A set \(S\) of \(n\) (opaque) line segments in the plane.

[Output] A set \(EL(S) = (EL_1(S), EL_2(S), \ldots, EL_k(S))\) \((1 \leq k \leq n)\) of the envelope layers of \(S\), where \(k\) is the number of layers, and \(EL_i(S)\) \((1 \leq i \leq k)\) is the \(i\)th envelope layer of \(S\).

**(Phase 1)** Use a segment tree, denoted as \(ST(S)\), which has \(m = O(\log n)\) levels, to cut the \(n\) segments of \(S\) into \(O(n \log n)\) subsegments and divide them into \(m\) separate groups, say \(G_1, G_2, \ldots, G_m\), where group \(G_i\) \((1 \leq i \leq m)\) corresponds to the subsegments associated on level \(i\) of \(ST(S)\). Reduce the envelope layers problem of \(G_i\) to the convex layers problem of \(|G_i|\) points.
(Phase 2) For each $i$ ($1 \leq i \leq m$) in parallel, find $EL(G_i)$ the envelope layers of $G_i$ by using our algorithm for the convex layers.

(Phase 3) Find the envelope of $EL(G_1), \ldots, EL(G_m)$ as follows.

1. Cutting $EL(G_1), \ldots, EL(G_m)$, by $n^{\frac{1-\epsilon}{2}}$ vertical lines into $n^{\frac{1-\epsilon}{2}}$ separate groups $H_1, H_2, \ldots, H_n^{\frac{1-\epsilon}{2}}$. Then, for each $i$ ($1 \leq i \leq n^{\frac{1-\epsilon}{2}}$) in parallel, compute the envelope layers of $H_i$, $EL(H_i) = (EL_1(H_i), EL_2(H_i), \ldots, EL_{h_i}(H_i))$, where $h_i$ is the number of layers in $H_i$, by Hershberger’s sequential algorithm [6].

2. Let $k = \max(h_1, h_2, \ldots, h_n^{\frac{1-\epsilon}{2}})$. Obtain $EL_t(S)$ ($1 \leq t \leq k$), i.e. the $t$th envelope layer of $S$, by concatenating $EL_t(H_1), EL_t(H_2), \ldots, EL_t(H_n^{\frac{1-\epsilon}{2}})$ ($EL_t(H_i)$ is empty if $t > h_i$).

Phase 1 can be done in $O(\log n)$ time using $n$ processors. When $1 \leq k \leq n^{\frac{2}{\epsilon}}$ ($0 < \epsilon < 1$), Phase 2 can be done in $O\left(\frac{n\log^2 n}{p}\right)$ time using $p$ processors, and Phase 3 can be done in $O\left(\frac{na(n)\log^3 n}{p}\right)$ time using $p$ processors, where $1 \leq p \leq n^{\frac{1-\epsilon}{2}}$. The proof can be found in [2].

References

Enhanced Sequence Reconstruction with DNA Microarray Application*

Extended Abstract

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Abstract. DNA sequencing by hybridization is potentially a powerful alternative to standard gel electrophoresis techniques. An important aspect of the approach is the design of the probing scheme and of the associated sequence reconstruction algorithm. Recently a novel probing scheme, whose performance is within a constant factor of the information theory bound, has settled the issue of asymptotic optimality. Thus, the research focus has shifted to the fine tuning of actual performance, with enormous potential for the life sciences. In this paper we discuss a new algorithmic device, called voting upon failure, which, exploiting the knowledge acquired in the course of the sequence reconstruction process, achieves typically a 20% performance improvement over the previous best technique, and comes at 90%-confidence within a factor 0.5 of the information-theory bound.

1 Introduction

A central application in molecular biology is the sequencing of DNA, i.e., the determination of the sequence of nucleotides of a chosen (fragment of a) DNA molecule. In recent years, a radically new technique has been proposed as an alternative to the traditional sequencing by gel electrophoresis. This technique, called Sequencing by Hybridization (SBH), proposed independently by several different research teams [BS91, D89, L88, P89, P91] is based on the use of a chip fabricated on a glass substrate. The active area of the chip is structured as a matrix, in each region of which (a feature) numerous copies of a specific oligonucleotide (a short sequence of DNA nucleotides) are implanted. The chip is immersed under controlled conditions within a solution of suitably labeled target DNA sequence. A copy of the target DNA will bind (hybridize) to an oligonucleotide if the oligonucleotide is complementary, in the Watson-Crick sense, to one of its subsequences. The labeling of the target allows visualization of the chip features containing binding oligonucleotides, thereby yielding a method for automatically probing the target sequence for specific subsequences.

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In summary, Sequencing by Hybridization consists of two fundamental steps. The first, biochemical in nature, is the acquisition, by complementary hybridization with a complete library of probes, of all subsequences (of a selected pattern) of a given unknown target sequence; the set of such subsequences is called the sequence *spectrum*. The second step, combinatorial in nature, is the algorithmic reconstruction of the sequence from its spectrum. Most traditional schemes were restricted to probes in the form of strings of *k* symbols (*k*-mers), which could achieve no better performance (expressed as the length of reliably reconstructible sequences) than the square root of the information-theoretic bound \[DF94, PF99\], as prominently noted in a review paper by well-known biologist Ed Southern \[S96\].

Recently a novel approach has been proposed \[PFU99, PU00\], whose performance is within a small factor of the optimum. The central feature of this approach is the adoption of a probing pattern where natural nucleotides are separated by gaps of “wild-cards”, intended to hybridize non-specifically with any natural base and proposed to be realized with artificial universal bases \[LB94\], i.e., bases that stack without binding.

It must be stressed that, since the question of the asymptotic performance has been settled, the focus of research in SBH sequence reconstruction has shifted to the refinement of the algorithmic procedures, in order to fine-tune computational methods which have enormous applicational potential in the life sciences. In this paper we shall briefly review the combinatorics underlying the approach, the known algorithmic technique, and illustrate a new property which enables a significant enhancement of the achievable performance.

## 2 Review of the Gapped-Probe Approach

A key feature of SBH is the adopted probing pattern, i.e., the structure of the spectrum probes. Notationally, a probing pattern is specified by a binary string (beginning and ending with a 1) where a 1 denotes a natural base and a 0 a universal base. For \(s + r = k\), an \((s, r)\) direct probing scheme has the pattern \(1^s(0^{s-1})^r\), while an \((s, r)\) reverse scheme has the pattern \((1^s)^r 1^s\). For notational convenience a probe is viewed as a string of length \((r + 1)s\) over the extended alphabet \(\mathcal{A} = \{ \text{A,C,G,T,}\ast\}\), where \(\ast\) denotes the “wild-card”.

In association with any such probing scheme the reconstruction of the sequence is algorithmically carried out symbol-by-symbol from one end to the other \[PFU99, PU00\]. The algorithmic primitive is the interrogation of the spectrum with a query of the form \(u\ast\), where \(u\) is a string of length \((r + 1)s - 1\) over \(\mathcal{A}\), which returns all matching probes. Performance is naturally measured as the length of sequences that are reconstructible with a given confidence level, under the standard hypothesis that the target sequence is generated by a maximum-entropy memoryless source (i.i.d. symbols).

The algorithm interrogates the spectrum with \(q\ast\), where \(q\) is the current \(((r + 1)s - 1)\)-symbol suffix of the reconstructed sequence. If only one probe is returned, the sequence is trivially extended; otherwise there is a potential ambiguity and
the algorithm attempts the extension of all potential paths up to a predetermined length $H$. The rationale is that the correct path is deterministically extended, whereas the extension of the spurious paths rests on the improbable presence of (fooling) probes in the spectrum. At the conclusion of this process, the sequence is extended with the common prefix of all surviving paths; otherwise, when such a prefix is empty, the existence of alternative paths of length $H$ is taken as evidence that their extension would continue with no definite bound, and reconstruction fails \cite{PU00}. The performance of such an algorithm at 90%-confidence comes within a factor slightly smaller than 0.5 of the theoretical optimum.

3 Analysis of the Failure Modes of the Reconstruction Algorithm

The best known algorithm for sequence reconstruction \cite{PU00} has two failure modes, distinguished on the basis of the paths issuing beyond the ambiguous branching embodying the reconstruction failure. The failure modes can be classified as follows:

1. Failure Mode 1. There are two paths identical except for their initial symbol (corresponding to the branching). This failure has been caused by $k$ fooling probes scattered, with possible overlaps, along the target sequence. On the basis of the spectrum alone there is no way to decide which of the multiple symbols detected at the branching corresponds to the correct extension.

2. Failure Mode 2. There are no two paths as specified for Failure Mode 1. If $H$ is large enough to make probabilistic extension of a spurious path to length $H$ extremely unlikely, then there must be an $((r+1)s - 1)$-segment of a spurious path, including or following the branching position which is identical to an actual segment occurring in the sequence. Obviously, in such a case extension of this path is deterministically assured and the reconstruction fails. (Such substrings are referred to as self-sustaining.) The self-sustaining segment agrees, entirely or partially, with an equally positioned segment of the correct path, the disagreements being concealed by fooling strings also occurring in the sequence.

Failure Mode 1 is well understood and discussed elsewhere \cite{PU00}. Suffice it here to recall that, denoting with $\alpha = 1 - e^{-m/4^k}$ the probability of occurrence of a specific probe in a target sequence of length $m$, an accurate estimate of the probability of failure by Mode 1 has the expression:

$$P_1 = 1 - e^{-3ma^k\left(1 + \frac{4^{r+1}}{s}\right)^r\left(1 + \frac{4^s}{ms}\right)^{s-1}}$$

Before analyzing Failure Mode 2, it is appropriate the study the two following examples (observed in simulations of random target sequences for the probing pattern (4,4): 1111 0001 0001 0001 0001). We have:
Example 1

```
GCAT TACT GAAC GGTG ACA [T] TGAT ATGA TGC...
[G] TCAC AAGC TAT...
```

```
... AC g a a c g g t t a c a g t c a c a a g C TAT ...
```

Example 2

```
ATAT GCAC GCCT [G] AGT GTAT ATTC TGGC CACT GCCA CG...
[T] AGT GTAT ATCA TAGG TACG TGAC AG...
```

```
... TCGC GCGG T AAC g t a t a t c a t a g g t a c g t g a C AG...
```

In the above examples the first two lines illustrate the branching event (symbols within brackets “[ ]”). Adequate prefixes of the two paths issuing from the branching are used to locate matching segments in the original sequence. The top path agrees with the putative sequence, whereas only a suffix of the bottom path agrees with an actual substring of the sequence. The latter is shown in the third line: the 19-symbol subsegment shown in lower case letters (referred to as the self-sustaining segment) supports the indefinite extension of the spurious path. This segment disagrees with the top (correct) segment in the positions labelled “∗”, and these disagreements must be concealed by fooling probes. Note that the self-sustaining segment may not begin more than 19 = (r + 1)s − 1 positions to the left of the branching; in Example 1 it starts on the left of the branching, in Example 2 on its right.

Next we analyze in detail this failure mode. The self-sustaining segment has length \( \nu = (r + 1)s - 1 \). This length is indeed sufficient for indefinite path extension. If the branching position is conventionally denoted as 0, the first position \( J \) of the self-sustaining segment, referred to as the offset, must satisfy the condition \( J \geq -\nu \). The failure events corresponding to each value \( J \) of the offset can be constructed by first selecting two positions in the sequence (in \( \binom{m}{2} \approx m^2/2 \) ways), so that its probability, for some coefficient \( \pi_J \), can be expressed as

\[
\frac{m^2}{2} \pi_J
\]

The term \( \pi_J \), referred to as the multiplier, depends upon the disagreements between the self-sustaining segment and the correct path.

The analysis of \( \pi_J \) is greatly aided by an \( s \)-row arrays (shown in Figure 1 for \( s = 5 \) and \( r = 3 \)), whose cells correspond to sequence positions with reference to the branching position and are numbered starting from \( -\nu = -(r + 1)s + 1 \) in column-major order. In the array we identify as the critical set the set of cells 0, 2s, \ldots, rs, rs + 1, rs + 2, \ldots, (r + 1)s − 1 (shown crossed-out in Figure 1),
Fig. 1. Illustration of offset domains for $J = -7$ and $J = 7$. Evidence is given to constrained symbols (light shading), critical set (crossed), and selectable rows (heavier shading).

which represent the positions of the last symbols of the fooling probes necessary to support the branching. A value $J$ of offset identifies in this array the domain (offset domain) of the self-sustaining segment (two typical offset domains are shown in Figure 1 for $J = -7$ and $J = 7$). For each value of the offset the following parameters are significant:

1. The number of deterministically constrained sequence symbols at the self-sustaining segment site. Since no fooling probe can conceal disagreements that precede the branching, all the symbols preceding the branching are constrained (cells shown lightly shaded in Figure 1(a)).
2. The size of the intersection with the critical set. Each cell of this intersection corresponds to a fooling probe which is to match the self-sustaining segment. (Such sets, shown horizontally hatched, have size 3 in 1(a) and size 6 in 1(b).)

3. The number of cells of the critical set with value smaller than $J$ (such a set is empty for $J \leq 0$ and is shown vertically hatched in Figure 1(b)). Each such cell identifies a weak fooling probe.

All cells in the self-sustaining segment domain with positive index and not belonging to the critical set are called selectable and may be sites of disagreements between the correct path and the self-sustaining segment (shown heavily shaded in Figure 1). The usefulness of the above diagram derives from the fact that, due to the pattern of fooling probes required by a disagreement (refer to the critical set as an example), rows of selectable cells are independent and contribute multiplicatively to $\pi_J$.

It follows that the basic components of the analysis are the rows of selectable cells. Consider a row of $u$ selectable cells ($u$-cell row). If there are no disagreements, then the row contributes $(1/4)^u$ to the multiplier; otherwise, let $h$ ($h = 1, \ldots, u$) be the leftmost cell of the row where a disagreement occurs. In such a case the contribution of the row is $(1/4)^{h-1}(3/4)^{u-h+1}$, since $h - 1$ symbols agree (each with probability $1/4$), the $h$-th symbol disagrees (with probability $3/4$) and must be confirmed by $u - h + 1$ fooling probes, regardless of the identity of the symbols in cells to its right on the row. Summing over $h$ we obtain

$$\frac{1}{4^u} + \sum_{h=1}^{u} \frac{1}{4^{h-1}} \frac{3}{4}^{u-h+1} = \frac{1}{4^u} (1 + 3\alpha \sum_{i=0}^{u-1} (4\alpha)^i)$$

We let for simplicity of notation $\beta_u = 1 + 3\alpha \sum_{i=0}^{u-1} (4\alpha)^i$, so that the contribution of a $u$-cell row becomes

$$\frac{\beta_u}{4^u}$$

For convenience of analysis, values of $J$ are aggregated on the basis of the size $\sigma$ of the intersection of the placement domain with the critical set. These sets, denoted Group I, ..., Group VI are illustrated for our running example in Figure 2, as differently shaded domains of cells. Cells in the same domain are covered by identical analysis.

In general, these sets of offset are:

1. Group I: $J = -\nu$, $\sigma = 0$;
2. Group II: $J = -\nu + 1 + is + j$, $i = 0, 1, \ldots, r - 1$, $j = 0, 1, \ldots, s - 1$;
3. Group III: $J = -s + 2$, $s + 3, \ldots, 0$, $\sigma = r + 1, r + 2, \ldots, k - 1$;
4. Group IV: $J = 1 + is + j$, $i = 0, 1, \ldots, r - 1$, $j = 0, 1, \ldots, s - 1$;
5. Group V: $J = rs + 1, rs + 2, \ldots, \nu$, $\sigma = s - 1, s - 2, \ldots, 1$;
6. Group VI: $J > \nu, \sigma = 0$.

In the full paper we shall report the detailed analysis of the contribution of each of these sets to the probability of failure. In this extended abstract, we illustrate in total detail the analysis of Groups I and II. It must be mentioned that, typically, the contribution of Group II accounts for about 4/5 of the total.
1. Group I: $J = -\nu, \sigma = 0$; The self-sustaining segment fully agrees with the correct path, and the corresponding event has multiplier $\pi_{-\nu} = 3/4^{\nu+1}$, because $\nu$ symbols are fully constrained and one (the branching) can be chosen in 3 ways.

2. Group II: $J = -\nu + 1 + is + j, i = 0, 1, \ldots, r - 1, j = 0, 1, \ldots, s - 1$, so that $\sigma = i + 1$. We observe that the $|J|$-prefix of the segment fully agrees with the correct segment, thereby constraining $|J|$ sequence symbols. (In fact, disagreements may only occur at selectable cells with index $\geq 0$.) For each value of $J$, we form the vector of the lengths of $s$ selectable-cell rows. This is illustrated in Table I, whose columns, indexed by the parameter $j$, are the length vectors.

<table>
<thead>
<tr>
<th>$j$</th>
<th>1</th>
<th>2</th>
<th>3</th>
<th>...</th>
<th>$s$</th>
</tr>
</thead>
<tbody>
<tr>
<td>row 1</td>
<td>0</td>
<td>0</td>
<td>0</td>
<td>...</td>
<td>0</td>
</tr>
<tr>
<td>row 2</td>
<td>$i$</td>
<td>$i+1$</td>
<td>$i+1$</td>
<td>...</td>
<td>$i+1$</td>
</tr>
<tr>
<td>row 3</td>
<td>$i$</td>
<td>$i$</td>
<td>$i+1$</td>
<td>...</td>
<td>$i+1$</td>
</tr>
<tr>
<td>...</td>
<td>...</td>
<td>...</td>
<td>...</td>
<td>...</td>
<td>...</td>
</tr>
<tr>
<td>row $s$</td>
<td>$i$</td>
<td>$i$</td>
<td>$i$</td>
<td>...</td>
<td>$i+1$</td>
</tr>
</tbody>
</table>

Placement $J = -\nu + is + j$ constrains $|J|$ sequence symbols and is accompanied by $(i+1)$ fooling probes of the critical set. Also, since the multiplicative contribution of a $u$-cell row is $\beta_u/4^u$, we conclude that the multiplier is

$$
\pi_J = \frac{3}{4} \frac{\alpha^{i+1}}{4^{i+j}} \left( \frac{\beta_{i+1}}{4^i+1} \right)^{j-1} \left( \frac{\beta_i}{4^i} \right)^{s-j} = \frac{3\alpha(4\alpha)^i}{4^\nu} \left( \frac{\beta_{i+1}}{\beta_i} \right)^{j-1} \beta_{s-1}^j
$$
If we sum over $j$ and $i$ we obtain the global contribution of Group II:

$$\sum_{i=0}^{r-1} \sum_{j=1}^{s} \frac{3\alpha(4\alpha)^i}{4^\nu} \left( \frac{\beta_{i+1}}{\beta_i} \right)^{j-1} \beta_{i}^{s-1} = \frac{1}{4^\nu} (\beta_r^s - 1)$$

4 Sequence Extension Based on Voting

In the presence of the any of the previously described failures, the standard algorithm is unable to provide a reliable reconstruction. Although for Failure Mode 2 prior knowledge of the approximate sequence length may be used to discriminate between the two paths being deterministically extended (by observing that the spurious extension will terminate with a sequence of length different from the nominal one), such a criterion is inapplicable to Failure Mode 1, since both extended paths are identical beyond the branching symbol.

However, there is one important item of information that the advanced algorithm fails to use, i.e., the content, when failure occurs, of the prefix $a_{(i,\ell)}$ of the putative sequence currently constructed. In such a situation we have two alternative sets of fooling probes, respectively corresponding to the two competing extensions. We may verify which of the two sets is better represented in $a_{(i,\ell)}$ and use this information to discriminate between the two alternatives. Following intuition, since failure has occurred, one set of fooling probes (the one causing the spurious extension) is guaranteed to exist, and its members are assumed to be scattered along the sequence. Therefore we adopt the criterion to reject as spurious the path whose fooling probes are better represented in the prefix $a_{(i,\ell)}$. Failure occurs also in the case of a tie, for in such a case discrimination is not possible. We now wish to analyze probabilistically the behavior of this policy, referred to as voting.

Assuming that there are just two alternative extensions (the event of more simultaneous extensions is extremely rare), the discrimination is carried out on the basis of the competing sets of fooling probes. The analysis is straightforward for Failure Mode 1, since in such a case each probe set consists of the $k$ probes sampling the branching position, because all other probes extend both paths. For Failure Mode 2, the analysis is more subtle. Referring to the preceding Example 2, each of the positions marked with an asterisk “∗” entails a train of at most $(r+1)$ fooling probes whose last symbol occurs not beyond the end of the self-sustaining segment. The union of such trains of probes is the fooling probe set for the chosen alternative. However, the self-sustaining segment cannot be determined by the algorithm; based on the fact that long transition segments are improbable, the algorithm constructs the discriminating sets on the basis of the disagreements occurring within $(r+1)s - 1$ positions from the branching. In the analysis, we shall use exactly $(r+1)s - 1$ probes, which is a valid criterion since each unrequired probe detected in $a_{(i,\ell)}$ adds 1 to both probe counts.

Let $1 < \ell < m$ be the position of failure and let $\mu_i$ be the common size of the competing probe sets, with $i = 1,2$ corresponding to the two modes (i.e., $\mu_1 = k$ and $\mu_2 = (r+1)s - 1$). Assume that there are $j$ probes of the incorrect
extension in the range \([1, \ell]\). Since failure has occurred, the \(\mu_i\) probes of the incorrect extension set are scattered along the sequence, and the probability that a specific probe of this set occurs in \(a_{(1, \ell)}\) is \(\gamma = \ell / m\). It follows that the probability of finding exactly \(j\) probes in \(a_{(1, \ell)}\) is

\[
p_j^{(i)}(\ell) = \binom{\mu_i}{j} \gamma^j (1 - \gamma)^{\mu_i - j}
\]

Let \(q_j^{(i)}(\ell)\) be the probability that there are at least \(j\) probes of the correct extension set in \(a_{(1, \ell)}\). Since the probability of occurrence of a specific probe is now \(\delta = 1 - e^{-\ell / 4^k}\), we obtain

\[
q_j^{(i)}(\ell) = \sum_{h=j}^{\mu_i} \binom{\mu_i}{h} \delta^h (1 - \delta)^{\mu_i - h}
\]

Therefore the quantity \(\sum_{j=0}^{\mu_i} p_j^{(i)}(\ell)q_j^{(i)}(\ell)\) is the probability of incorrect vote in correspondence to a failure of Mode \(i\). We also let \(\pi_i(\ell)\) denote \(\text{Prob}(\ell)\) for Failure Mode \(i\). We note that \(\pi_1(\ell)\) is uniformly distributed in the range \([1, m]\), whereas \(\pi_2(\ell)\) decreases linearly in the same range (in fact, in this case \(\ell\) is the smaller of two values chosen in \([1, m]\)). The probability that voting fails, conditional on the occurrence of the appropriate type of failure, is therefore

\[
P(\text{vote failure}|\text{Mode } i) = \sum_{\ell=1}^{m} \pi_i(\ell) \sum_{j=0}^{\mu_i} p_j^{(i)}(\ell)q_j^{(i)}(\ell)
\]

\[
= \sum_{\ell=1}^{m} \pi_i(\ell) \sum_{j=0}^{\mu_i} \binom{\mu_i}{j} \gamma^j (1 - \gamma)^{\mu_i - j} \sum_{h=j}^{\mu_i} \binom{\mu_i}{h} \delta^h (1 - \delta)^{\mu_i - h}
\]

\[
\tag{1}
\]

In conclusion we obtain the probability of failure \(\phi(m)\) as a function of \(m\)

\[
\phi(m) = P(\text{vote failure}|\text{Mode 1})P_1 + P(\text{vote failure}|\text{Mode 2})P_2
\]

\[
< P_1 + P_2
\]

and the equation \(\phi(m*) = \epsilon\) quantifies, for a given \(\epsilon\), the performance enhancement due to the voting approach.

Unfortunately, Expression (1) is not easily amenable to a closed form. However, numerical calculations of \(\phi(m)\) have been carried out and are summarized in Figure 3 for a specific illustration: A \((4, 4)\) probing scheme, i.e., \(k = 8\). Although specific, this example is quite significant because it represents, in a sense, the accepted state of the art in terms of the microarray size. Figure 3 displays the probability of incorrect voting for Mode 1 failure as a function of \(\ell\) (point of failure) and \(m\) (sequence length).

As expected, this function increases with \(m\) for fixed \(\ell\) and tends to 1 as \(\ell\) decreases, since in such a condition the tie \(j = h = 0\) becomes the most
Fig. 3. Probability of incorrect voting for Mode 1 failure as a function of \( \ell \) and \( m (\ell \leq m) \).

likely event. The quantities \( P(\text{vote failure}|\text{Mode } i) \) are obtained by averaging the previous functions over the parameter \( \ell \). Finally, the quantities \( P_1, P_2, P(\text{vote failure}|\text{Mode 1}) \) and \( P(\text{vote failure}|\text{Mode 2}) \), each a function of \( m \), are combined as in (2) to obtain the desired performance function \( \phi(m) \). The latter is presented as \( 1 - \phi(m) \) in Figure 4, to be contrasted with the probability of success of the previous algorithm [PU00] without the voting provision.

Shown in the same figure are sets of experimental data points for the two situations, which are in remarkable agreement with the corresponding analyses. At 90%-confidence the length of the correctly reconstructible sequence increases by nearly 20%, which quantifies the enhancement afforded by the approach discussed in this paper, quite likely at the frontier of achievable performance.

References


Fig. 4. Comparison of the algorithms without (left) and with (right) voting provision for a (4, 4) probing scheme (probability of success versus $m$). Analytical curves for the two cases are paired with corresponding extensive simulation data.


Non-approximability
of Weighted Multiple Sequence Alignment

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Abstract. We consider a weighted generalization of multiple sequence alignment with sum-of-pair score. Multiple sequence alignment without weights is known to be \(\mathcal{NP}\)-complete and can be approximated within a constant factor, but it is unknown whether it has a polynomial time approximation scheme. Weighted multiple sequence alignment can be approximated within a factor of \(O(\log^2 n)\) where \(n\) is the number of sequences.

We prove that weighted multiple sequence alignment is \(\text{MAX } \mathcal{SNP}\)-hard and establish a numerical lower bound on its approximability, namely \(\frac{324}{323} - \epsilon\). This lower bound is obtained already for the simple binary weighted case where the weights are restricted to 0 and 1. Furthermore, we show that weighted multiple sequence alignment and its restriction to binary weights can be approximated exactly to the same degree.

1 Introduction

Multiple sequence alignment (MSA) is an important problem in computational biology. The alignment of a group of protein or nucleotide sequences yields information about the relationships between these sequences and it is also used to detect similarities (so called “homologous regions”) between them. This information is applied in constructing evolutionary trees and finding coherences between the function and structure of proteins and their sequences.

Many objective functions have been suggested to measure the quality of a multiple sequence alignment. One of the most widely used is the so called sum-of-pair score (SP-score, Carrillo et al. [6]).

MSA with SP-score is known to be \(\mathcal{NP}\)-complete (Wang et al. [12]). For the case that the scoring function does not have to be a metric, Just has shown that MSA with SP-score is \(\text{MAX } \mathcal{SNP}\)-hard [9]. Akutsu et al. have investigated the multiple sequence alignment problem under several scoring functions, namely \(\#LOG\#\)-score and \(IC\)-score [1]. They have shown that a variant of the multiple sequence alignment problem called local multiple alignment is \(\text{MAX } \mathcal{SNP}\)-hard under these scoring schemes.

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However, if the scoring function fulfills the triangle inequality, no lower bound for this problem is known so far. The complexity of MSA over an alphabet of fixed size with metric SP-scoring functions is of main interest. According to Jiang et al. the approximability of MSA with metric SP-score is an important open problem in computational biology \[8\].

To represent existing knowledge about the relationships of the sequences considered, a weighted variant of MSA was introduced by Wu et al. \[13\]. Each pair of sequences is assigned a nonnegative value reflecting their degree of relationship. This means that a pair which is assumed to be closely related will be assigned a high weight while a less related pair will be assigned a smaller weight. This generalization of MSA is called weighted MSA, or WMSA for short.

In this paper we also examine a restricted version of WMSA called binary weighted MSA (BMSA), where the weights are restricted to 0 and 1. The binary weights can be used to represent an arbitrary graph over which multiple sequence alignments can be determined. We will prove that BMSA is equivalent to WMSA with respect to their approximability. Thus, an approximation algorithm for BMSA directly yields an approximation algorithm for the general case with the same performance ratio. Moreover, we prove the MAX $SNP$-hardness and a numerical lower bound for the approximability of BMSA. These results are obtained even if the sequences are of fixed length and the alphabet is of fixed size. Thus, the difficulty of multiple sequence alignment is caused by the number of sequences, not by their length.

In the next section we give a formal definition of the problems considered. The reduction from WMSA to BMSA is presented in section 3. In section 4 we prove a lower bound for the approximability of a problem called MAX-E2-neg-Lin2. This result will be used in section 5 to prove a lower bound for the approximability of BMSA.

## 2 Definitions and Notations

Let $\Sigma$ be an alphabet and $\Sigma' := \Sigma \cup \{-\}$, where “−” denotes a gap symbol. $S[l]$ denotes the $l$-th symbol of a sequence $S$. Let $\mathcal{S} = \{S_1, \ldots, S_n\}$ be a family (a multiset) of sequences over $\Sigma$. An alignment of $\mathcal{S}$ is a family $\mathcal{A} = \{\tilde{S}_1, \ldots, \tilde{S}_n\}$ of sequences over $\Sigma'$ such that all $\tilde{S}_i$ have equal length and $\tilde{S}_i$ is obtained from $S_i$ by inserting gaps. The following is an example of an alignment of three sequences ATTCTG, TTCTTTG and ATTGTT.

\[
\begin{align*}
ATTCT & \quad \quad -G \\
-\quad TTCTTTG \\
ATTGT & \quad \quad T- 
\end{align*}
\]

A function $d : \Sigma'^2 \rightarrow \mathbb{N}$ will be called scoring function if it is a metric, i.e. for any $x, y, z \in \Sigma'$ we have $d(x, y) = 0$ iff $x = y$, $d(x, y) = d(y, x)$, and $d(x, z) \leq d(x, y) + d(y, z)$. We define the distance of two sequences $\tilde{S}_i$ and $\tilde{S}_j$ of length $l$ as $D(\tilde{S}_i, \tilde{S}_j) := \sum_{k=1}^l d(\tilde{S}_i[k], \tilde{S}_j[k])$. 

Carrillo and Lipman introduced a scoring scheme for alignments called sum-of-pair score (SP-score, [6]). The SP-score of an alignment \( A = \{ \tilde{S}_1, \ldots, \tilde{S}_n \} \) is defined by \( D(A) := \sum_{1 \leq i < j \leq n} D(\tilde{S}_i, \tilde{S}_j) \). Multiple sequence alignment (MSA) is the problem of finding an alignment with minimum SP-score.

Wu et al. generalized MSA to weighted sum-of-pair score [13]. The weights are given by \( W := (w_{S_i, S_j}) \) for \( S_i, S_j \in S \), a symmetric matrix of nonnegative integers. Then the weighted SP-score of an alignment \( A \) is \( D_W(A) := \sum_{1 \leq i < j \leq n} w_{S_i, S_j} \cdot D(\tilde{S}_i, \tilde{S}_j) \). This generalization is called weighted multiple sequence alignment (WMSA). The aim is to find an alignment with minimum weighted SP-score.

An instance of WMSA is a 4-tuple \( (\Sigma, S, d, W) \). We consider the case of a fixed alphabet \( \Sigma \) and a fixed scoring function \( d \). Thus, a problem instance of WMSA is given by a pair \( (S, W) \). It is easy to see that any lower bound for this case also holds if we allow arbitrary scoring functions and alphabets.

A special case of WMSA is binary weighted MSA (BMSA), where the weights are restricted to 0 and 1.

It has been shown that MSA with SP-score is \( \mathcal{NP} \)-complete [12]. For an arbitrary fixed constant \( r \), MSA can be approximated in polynomial time within a factor of \( 2 - \frac{r}{n} \), where \( n \geq r \) is the number of sequences [4]. It is unknown whether MSA admits a polynomial time approximation scheme (PTAS, see e.g. Ausiello et al. [3]). WMSA with arbitrary weights can be approximated within a factor of \( O(\log^2 n) \) [13]. Using a technique of Bartal [5] one can obtain a randomized \( O(\log n \cdot \log \log n) \) approximation.

Papadimitriou et al. introduced a class of optimization problems called MAX \( \mathcal{SNP} \) [10]. They showed that there exist problems which are MAX \( \mathcal{SNP} \)-complete with respect to L-reductions. In the following, \( \text{opt}(I) \) denotes the optimal score of an instance \( I \) of an optimization problem. For example, \( \text{opt}(S) \) denotes the score of an optimal (weighted) alignment of \( S \).

**Definition 1.** Let \( \Pi \) and \( \Pi' \) be two optimization problems. Then \( \Pi \) L-reduces to \( \Pi' \) if there exist polynomial time computable functions \( f_1, f_2 \) and constants \( \gamma_1, \gamma_2 > 0 \) such that for each instance \( I \) of \( \Pi' \):

1. Function \( f_1 \) produces an instance \( I' = f_1(I) \) of \( \Pi' \) such that \( \text{opt}(I') \leq \gamma_1 \cdot \text{opt}(I) \).
2. Given a solution \( S' \) of \( I' \) with cost \( c'(S') \), function \( f_2 \) produces a solution \( S = f_2(I, S') \) of \( I \) with cost \( c(S) \) such that \( |c(S) - \text{opt}(I)| \leq \gamma_2 \cdot |c'(S') - \text{opt}(I')| \).

No MAX \( \mathcal{SNP} \)-hard problem has a PTAS, unless \( \mathcal{NP} = \mathcal{P} \) (Arora et al. [2]).

**3 Reduction from WMSA to BMSA**

Let \( S = \{ S_1, \ldots, S_n \} \) be a family of sequences over \( \Sigma \) and \( W = (w_{S_i, S_j}) \) be a weight matrix. Let \( l \) be the maximal length of the sequences in \( S \) and \( d_{\text{max}} \) be the maximum of the scoring function \( d \). We assume that the weights and the scoring function are unary coded. This does not seem to be a restriction because in practice the weights are very small and the scoring function is fixed.
We construct a family of sequences $S'$ as an instance of BMSA as follows. Let $K := 2 \cdot d_{\text{max}} \cdot l$. For a sequence $S_j \in S$ generate $K$ copies $T_j^k \in S'$ ($1 \leq k \leq K$) of this sequence. Furthermore, for each $1 \leq i \leq n$ construct $w_{S_i, S_j}$ copies $S_j^{i, \mu} \in S'$ ($1 \leq \mu \leq w_{S_i, S_j}$) of $S_j$. The weight matrix $W' = (w'_{i, j})_{1 \leq i, j \leq S'}$ is given by

$$w'_{i, j} := \begin{cases} 1 & \text{if } I \equiv S_j^{i, \mu} \text{ and } J \equiv S_j^{i, \mu}, \\ 1 & \text{if } I \equiv S_j^{i, \mu} \text{ and } J \equiv T_j^k \text{ or vice versa}, \\ 0 & \text{otherwise,} \end{cases}$$

where $A \equiv B$ means that $A$ and $B$ are not only equal but denote the same sequence.

Since the weights and the scoring function are unary coded, the input size $N$ of the instance of WMSA fulfils the bound $N \in \Omega(n \cdot l + \sum_{i, j=1}^n w_{S_i, S_j})$. On the other hand, the input size $N'$ of the constructed instance of BMSA satisfies

$$N' \in O(n \cdot K \cdot l + l \cdot \sum_{i, j=1}^n w_{S_i, S_j} + (n \cdot K + \sum_{i, j=1}^n w_{S_i, S_j})^2).$$

Note that $N'$ is polynomially bounded by $N$.

**Lemma 1.** If $S$ has an alignment $A$ with weighted score $D_W(A)$ then $S'$ has an alignment $A'$ with score $D_W(A') = D_W(A)$.

**Proof.** Let $A = \{\tilde{S}_1, \ldots, \tilde{S}_n\}$ be an alignment of $S$ with weighted score $D$. We obtain an alignment $A' = \{\tilde{A} | A \in S'\}$ of $S'$ by setting $\tilde{T}_j^k = \tilde{S}_j$ and $\tilde{S}_j^{i, \mu} = \tilde{S}_j$ for all $j, k, i, \mu$. The score of $A'$ with respect to the weight matrix $W'$ is

$$D_W(A') = \sum_{i, j=1}^n \sum_{\mu=1}^K \sum_{k=1}^K D(\tilde{S}_j^{i, \mu}, \tilde{T}_j^k) + \sum_{1 \leq i < j \leq n} \sum_{\mu=1}^n D(\tilde{S}_i^{j, \mu}, \tilde{S}_j^{i, \mu}) = D_W(A).$$

\[ \square \]

**Lemma 2.** Given an alignment $A'$ of $S'$ with weighted score $D_W(A')$ we can construct an alignment $A$ of $S$ with less or equal score in polynomial time.

**Proof.** Let $A' = \{\tilde{A} | A \in S'\}$ be an arbitrary alignment of $S'$ with score $D_W(A')$. The copies of a sequence $S_j \in S$ will be called consistent if there exists a sequence $B_j$ with $\tilde{T}_j^k = B_j$ and $\tilde{S}_j^{i, \mu} = B_j$ for all $k, i, \mu$. The sequence $B_j$ is called block.

We consider the case that for some $j_0$ the copies of $S_{j_0}$ are not consistent and distinguish two cases. First, if not all $\tilde{T}_j^k$ are equal, let

$$D_k := \sum_{i=1}^n \sum_{\mu=1} w_{S_i, S_j} D(\tilde{T}_j^k, \tilde{S}_j^{i, \mu})$$

be the score of $\tilde{T}_j^k$ with the sequences $\tilde{S}_j^{i, \mu}$. Choose $k_0$ such that $D_{k_0}$ is minimal among all $D_k$ and set $\tilde{T}_{j_0}^k = \tilde{T}_{j_0}^{k_0}$ for all $k \neq k_0$. This way we obtain a new alignment with less or equal score.
Now we consider the case that there exists a $B_{j_0}$ such that $\tilde{T}^k_{j_0} = B_{j_0}$ for all $k$. Then there exists a sequence $\tilde{S}^1_{j_0, \mu_0} \neq B_{j_0}$. This sequence yields at least score $K$ with the sequences $\tilde{T}^k_{j_0}$, because it yields a score of at least 1 with every $\tilde{T}^k_{j_0}$. Set $S^1_{j_0, \mu_0} = B_{j_0}$. Then $\tilde{S}^1_{j_0, \mu_0}$ yields score 0 with any $\tilde{T}^k_{j_0}$ and at most score $K$ with $S^1_{j_0, \mu_0}$. Thus, the new alignment has less or equal score.

By these modifications we iteratively obtain a new alignment of $S'$ such that for any $j \in \{1, \ldots, n\}$ the copies of $S_j$ are consistent with block $B_j$. The blocks of $S'$ induce an alignment $A = \{B_1, \ldots, B_n\}$ of $S$ with score $D_W(A) = \sum_{1 \leq i < j \leq n} w_{S_i} s_j \cdot D(B_i, B_j) \leq D_W(A')$.

With these results we have shown that a $\lambda$-approximation for BMSA can be used as a $\lambda$-approximation for WMSA. Thus, the following theorem holds.

**Theorem 1.** If BMSA can be approximated within a constant factor $\lambda$ in polynomial time, then WMSA can also be approximated within $\lambda$ in polynomial time.

## 4 The Non-approximability of MAX-E2-neg-Lin2

We consider the multiplicative group $\{1, -1\}$. Let $G = \{G_1, \ldots, G_t\}$ be a multiset of linear equations over the variables $U = \{x_1, \ldots, x_r\}$, $G_i \equiv x_{\alpha_{i,1}} \cdots x_{\alpha_{i,k}} = a_i, k \geq 2, \alpha_{i,q} \in \{1, \ldots, r\}$, and $a_i \in \{1, -1\}$ is a constant. MAX-Ek-Lin2 is the optimization problem of finding the maximum number of simultaneously satisfiable equations. A restriction of MAX-Ek-Lin2 is MAX-Ek-neg-Lin2, where $a_i = -1$ for all $1 \leq i \leq t$.

MAX-E2-neg-Lin2 is exactly the problem MAX-Cut (see e.g. [3]) where the equations correspond to the edges, the variables correspond to the nodes, and multiple edges are allowed. Therefore, MAX-E2-neg-Lin2 is MAX $\mathcal{NP}$-complete [10]. We use MAX-E2-neg-Lin2 here due to the simpler notation.

An instance of MAX-Ek-Lin2 or MAX-Ek-neg-Lin2 consisting of $t$ equations will be called $\eta$-satisfiable iff $\eta \cdot t$ is the maximum number of simultaneously satisfiable equations. Håstad proved in [7] that it is $\mathcal{NP}$-hard to distinguish $(1 - \epsilon)$-satisfiable and $(1/2 + \epsilon)$-satisfiable instances of MAX-E3-Lin2 for any $\epsilon > 0$.

Instead of the known lower bound for the approximability of MAX-Cut (Håstad [7] and Trevisan et al. [11]) we will construct a reduction from MAX-E3-Lin2 to MAX-E2-neg-Lin2 to prove that it is $\mathcal{NP}$-hard to distinguish $(15/22 - \epsilon)$- and $(17/22 + \epsilon)$-satisfiable instances of MAX-E2-neg-Lin2 for any $\epsilon > 0$; the gadget used by Trevisan et al. [11] does not yield such a gap directly. This result will be used in section 5 to establish the lower bound for the approximability of BMSA.

We will now reduce MAX-E3-Lin2 to MAX-E2-neg-Lin2. Let $G = \{G_1, \ldots, G_t\}$ be a multiset of equations over variables $U$, $G_i \equiv x_{\alpha_{i,1}} \cdots x_{\alpha_{i,2}} x_{\alpha_{i,3}} = a_i$.

We construct an instance $G'$ of MAX-E2-neg-Lin2 with $22 \cdot t$ equations and $4 \cdot t + 2 \cdot r + 2$ variables. The reduction is similar to the reduction from MAX-E3-Lin2 to MAX-E2-Lin2 in [7]. The set of variables $U'$ is given by

$$U' = \{x_j^+, x_j^- | 1 \leq j \leq r\} \cup \{z^+, z^-\} \cup \{p_{i,1}, p_{i,2}, p_{i,3}, p_{i,z} | 1 \leq i \leq t\}.$$
Lemma 4. Given an arbitrary assignment for an equation of an instance of MAX-E2-neg-Lin2, then the negated assignment also satisfies the equation. So without loss of generality we assume that in any case $z^+ = 1$.

We interpret $x_j^+ = x_j$. We call an assignment consistent for $x_j$ if $x_j^+ \neq x_j^-$ and therefore $x_j^+ = x_j = (-x_j^-)$. An assignment that is consistent for every $x_j$ and where $z^+ \neq z^-$ is called consistent.

For an equation $G_i \triangleq x_{\alpha_{i,1}} \cdot x_{\alpha_{i,2}} \cdot x_{\alpha_{i,3}} = a_i$ we construct the twelve equations

\[
\begin{align*}
  x_{\alpha_{i,q}}^+ \cdot p_{i,q}^- &= -1 & \text{for } q, q' = 1, 2, 3 \text{ and } q \neq q', \\
  x_{\alpha_{i,q}}^- \cdot p_{i,z}^- &= -1 & \text{for } q = 1, 2, 3, \\
  x_{\alpha_{i,q}}^- \cdot p_{i,q}^- &= -1 & \text{for } q = 1, 2, 3.
\end{align*}
\]

We add either the four equations $z^+ \cdot p_{i,q}^- = -1 \ (q = 1, 2, 3)$ and $z^- \cdot p_{i,z}^- = -1$ if $a_i = 1$ or the four equations $z^- \cdot p_{i,q}^- = -1 \ (q = 1, 2, 3)$ and $z^+ \cdot p_{i,z}^- = -1$ if $a_i = -1$. For every equation in $G$ we construct the three equations $x_{\alpha_{i,q}}^+ \cdot x_{\alpha_{i,q}}^- = -1 \ (q = 1, 2, 3)$. Finally, we add the equation $z^+ \cdot z^- = -1$ three times. Note that $G'$ contains $3 \cdot t$ times the equation $z^+ \cdot z^- = -1$. Let $n_j$ be the number of occurrences of the variable $x_j$ in $G$. Then $G'$ contains $n_j$ times the equation $x_j^+ \cdot x_j^- = -1$.

For every equation $G_i \in G$ we have constructed 22 equations for $G'$. These 22 equations are called the representation of $G_i$.

Lemma 3. Let an arbitrary assignment for $U$ be given. Assign $z^- = -1$ and $x_j^+ = x_j$, $x_j^- = (-x_j)$ for $j = 1, \ldots, r$. Then for any $i \in \{1, \ldots, t\}$ there exists an assignment for $p_{i,1}, p_{i,2}, p_{i,3},$ and $p_{i,z}$ such that 18 equations of the representation of $G_i$ are satisfied if $G_i$ is satisfied by the given assignment and 16 equations of the representation are satisfied if $G_i$ is not satisfied.

It is not possible to satisfy more than 18 equations of the representation if $G_i$ is satisfied by the assignment and to satisfy more than 16 equations if $G_i$ is not satisfied by the assignment.

Proof. The lemma can be proved by testing all possible assignments. \qed

If an assignment for $U$ satisfies $g$ of the $t$ equations of $G$, then the corresponding consistent assignment for $U'$ satisfies $16 \cdot t + 2 \cdot g$ equations of $G'$. This assignment can be found efficiently by adjusting the assignment for $p_{i,1}, p_{i,2}, p_{i,3},$ and $p_{i,z}$. On the other hand, a consistent assignment for $U'$ that satisfies $16 \cdot t + 2 \cdot g$ equations of $G'$ yields an assignment for $U$ that satisfies $g$ equations of $G$.

Lemma 4. Given an arbitrary assignment for $U'$ that satisfies $16 \cdot t + 2 \cdot g$ equations of $G'$, a consistent assignment that satisfies at least this amount of equations of $G'$ can be computed in polynomial time.

Proof. First assume that $z^+ = z^-$ in the given assignment. Then the $3 \cdot t$ equations $z^+ \cdot z^- = -1$ are not satisfied by the assignment. Let $z^- = (-z^+)$. Then these $3 \cdot t$ equations will be satisfied. On the other hand, $z^-$ occurs in only $3 \cdot t$
other equations. Thus, at most $3 \cdot t$ equations are no longer satisfied. Altogether
the number of satisfied equations is not decreased by this modification.

If there exists a $j$ with $x_j^+ = x_j^-$, then there are $n_j$ equations $x_j^+ \cdot x_j^- = -1$ that are not satisfied by the assignment. Let $x_j^- = (-x_j^+)$. Then the $n_j$ equations $x_j^+ \cdot x_j^- = -1$ are satisfied by the modified assignment. On the other hand $x_j^-$ occurs in only $n_j$ other equations. Thus, at most $n_j$ equations are no longer satisfied. The number of satisfied equations is thus not decreased by this modification.

This way we iteratively obtain a consistent assignment. Obviously, the modifications can be computed in polynomial time.

Now we can prove the following theorem used in section 5.

**Theorem 2.** For any $\epsilon > 0$ it is $\mathcal{NP}$-hard to distinguish $(\frac{18}{22} - \epsilon)$- and $(\frac{17}{22} + \epsilon)$-satisfiable instances of MAX-E2-neg-Lin2.

**Proof.** An instance of MAX-E3-Lin2 is $\eta$-satisfiable iff the corresponding instance of MAX-E2-neg-Lin2 is $(\frac{16+2\cdot \eta}{22})$-satisfiable. According to Håstad [7] it is $\mathcal{NP}$-hard to distinguish $(1 - \xi)$- and $(\frac{1}{2} + \xi)$-satisfiable instances of MAX-E3-Lin2 for any $\xi > 0$. Thus, it is $\mathcal{NP}$-hard to distinguish $(\frac{16+2-(1-\xi)}{22})$- and $(\frac{16+2(\frac{1}{2}+\xi)}{22})$-satisfiable instances of MAX-E2-neg-Lin2. Choosing $\xi = 11 \cdot \epsilon$ completes the proof.

Since MAX-Cut and MAX-E2-neg-Lin2 are exactly the same problem, we obtain the same approximability gap for MAX-Cut.

**Corollary 1.** For any $\epsilon > 0$ it is $\mathcal{NP}$-hard to decide whether the maximum cut of an instance $G = (V, E)$ (where multiple edges are allowed) of MAX-Cut consists of at most $(\frac{17}{22} + \epsilon) \cdot |E|$ or at least $(\frac{18}{22} - \epsilon) \cdot |E|$ edges.

## 5 The Non-approximability of BMSA

In this section we reduce MAX-E2-neg-Lin2 to BMSA. Let $\mathcal{G} = \{G_1, \ldots, G_t\}$ be an instance of MAX-E2-neg-Lin2 over a set of variables $U = \{x_1, \ldots, x_r\}$, $G_i \equiv x_{\alpha_{i,1}} \cdot x_{\alpha_{i,2}} = -1$, $\alpha_{i,q} \in \{1, \ldots, r\}$. We construct a family of sequences $S = \{Z\} \cup \{X_j\}_{j = 1, \ldots, r} \cup \{Y_{i,1}, Y_{i,2}\}_{i = 1, \ldots, t}$ over the alphabet $\Sigma = \{\bullet, \circ, \times\}$. Let $Z := \circ\circ\circ\circ\circ\circ\circ\circ$ be a sequence of length 8. $Z$ will be used as a control sequence. For $j \in \{1, \ldots, r\}$ let $X_j := \bullet\circ\circ\circ\circ\circ\circ\circ\circ\circ$ be a sequence of length 9 that represents the variable $x_j \in U$. For each $i \in \{1, \ldots, t\}$ create two sequences $Y_{i,1} := \bullet\circ\circ\circ\circ\circ\circ\circ\circ\circ$ and $Y_{i,2} := \circ\circ\circ\circ\circ\circ\circ\circ\circ$ each of length 9. $Y_{i,q}$ represents the variable $x_{\alpha_{i,q}}$ in $G_i$.

The scoring function is given in the following table. Note that it is a metric.

<table>
<thead>
<tr>
<th></th>
<th>-</th>
<th>⊗</th>
<th>•</th>
<th>○</th>
<th>⊗</th>
<th>×</th>
</tr>
</thead>
<tbody>
<tr>
<td>-</td>
<td>0</td>
<td>1</td>
<td>2</td>
<td>5</td>
<td></td>
<td></td>
</tr>
<tr>
<td>⊗</td>
<td>1</td>
<td>0</td>
<td>1</td>
<td>4</td>
<td></td>
<td></td>
</tr>
<tr>
<td>○</td>
<td>2</td>
<td>1</td>
<td>0</td>
<td>3</td>
<td></td>
<td></td>
</tr>
<tr>
<td>⊗</td>
<td>5</td>
<td>4</td>
<td>3</td>
<td>0</td>
<td></td>
<td></td>
</tr>
</tbody>
</table>
The weight matrix $W = (w_{i,j})_{i,j \in S}$ is given by

$$w_{i,j} := \begin{cases} 
1 & \text{if } I \equiv Y_{i,q} \text{ and } J \equiv Y_{i,q'}, \\
1 & \text{if } I \equiv Z \text{ and } J \equiv Y_{i,q} \text{ or vice versa}, \\
1 & \text{if } I \equiv Y_{i,q} \text{ and } J \equiv X_{\alpha_{i,q}} \text{ or vice versa}, \\
0 & \text{otherwise}.
\end{cases}$$

The set $S_i = \{Y_{i,1}, Y_{i,2}, X_{\alpha_{i,1}}, X_{\alpha_{i,2}}\}$ will be called the representation of $G_i$. Note that in general a sequence $X_j$ occurs in more than one representation.

Let $A = \{\tilde{S}|S \in S\}$ be an alignment of $S$. Then $D_i(A)$ denotes the score of the equation $G_i$, $D_i(A) = D(\tilde{Y}_{i,1}, \tilde{Y}_{i,2}) + D(\tilde{Y}_{i,1}, \tilde{X}_{\alpha_{i,1}}) + D(\tilde{Y}_{i,2}, \tilde{X}_{\alpha_{i,2}}) + D(\tilde{Y}_{i,1}, \tilde{Z}) + D(\tilde{Y}_{i,2}, \tilde{Z})$. By the construction of the weight matrix, $D_W(A) = \sum_{i=1}^{t} D_i(A)$ holds.

**Definition 2.** An alignment $A = \{\tilde{S}|S \in S\}$ of $S$ will be called variable-consistent with respect to an assignment for $U$ if, after eliminating all columns consisting solely of gaps (which do not affect the score), the following holds for all $j$, $i$, and $q$:

1. $\tilde{Z} = -Z$.
2. $\tilde{X}_j = \begin{cases} X_j & \text{if } x_j = -1 \\
- X_j & \text{if } x_j = 1 \end{cases}$
3. $\tilde{Y}_{i,q} = \begin{cases} Y_{i,q} & \text{if } x_{\alpha_{i,q}} = -1 \\
- Y_{i,q} & \text{if } x_{\alpha_{i,q}} = 1 \end{cases}$

The following lemma follows immediately from this definition.

**Lemma 5.** An alignment is variable-consistent iff for all $i = 1, \ldots, t$ and $q = 1, 2$ the following properties hold:

A. Either $Y_{i,q}[1]$ or $Y_{i,q}[9]$ matches a gap in $Z$. No other character of $Z$ or $Y_{i,q}$ matches a gap in the other sequence.

B. No character in either of the two sequences $Y_{i,q}$, $X_{\alpha_{i,q}}$ matches a gap in the other sequence. \[ \square \]

These properties are referred to as property A and B. The following is an example of a variable-consistent alignment representing the equation $G_i \equiv x_1 \cdot x_2 = -1$ which is satisfied by $x_1 = -1$ and $x_2 = 1$.

$$\begin{align*}
\tilde{Y}_{i,1} & = \bullet \circ \circ \times \circ \times \circ \circ \bullet \circ \circ \\
\tilde{Y}_{i,2} & = -\bullet \circ \circ \times \circ \times \circ \circ \bullet \\
\tilde{X}_1 & = \bullet \circ \circ \circ \circ \circ \circ \circ \bullet \\
\tilde{X}_2 & = -\bullet \circ \circ \circ \circ \circ \circ \circ \bullet \\
\tilde{Z} & = -\circ \circ \circ \circ \circ \circ \circ \circ
\end{align*}$$

Note the functional region of a pair $Y_{i,1}, Y_{i,2}$ given by the triples $\times \circ \times$ and $\circ \times \circ$. If $Y_{i,1}$ and $Y_{i,2}$ represent the same value, the functional region yields a weighted score of 9. Otherwise, it yields a weighted score of 3. If an alignment $\mathcal{A}$ is variable-consistent, we have $D_i(\mathcal{A}) = 29$ if $G_i$ is satisfied by the represented assignment and $D_i(\mathcal{A}) = 31$ otherwise.
The next two lemmas have similar proofs. Thus, we only give a proof of the first.

**Lemma 6.** Alignments of the pairs \( \{Y_{i,1}, Z\} \) and \( \{Y_{i,2}, Z\} \) yield scores of 8 and 5, respectively, if they fulfil property A. Violating property A yields scores of at least 10 and 7, respectively.

*Proof.* An alignment of \( \{Y_{i,1}, Z\} \) that fulfils property A yields score 8.

Let us consider an alignment of \( \{Y_{i,1}, Z\} \) that does not fulfil property A. Then at least one of the characters \( Y_{i,1}[2], \ldots, Y_{i,1}[8], Z[1], \ldots, Z[8] \) matches a gap in the other sequence.

We distinguish two cases. If there is an “×” in \( Y_{i,1} \) matching a gap in \( Z \), then the alignment yields a score of 5 for this “×” plus 3 for the other “×” plus 1 for each “•”. So altogether it yields a score of at least 10.

On the other hand consider the case that no “×” in \( Y_{i,1} \) matches a gap in \( Z \).

Then there is a “◦” in \( Y_{i,1} \) or \( Z \) matching a gap in the other sequence. So the alignment yields a score of 3 for each “×” plus 1 for each “•” plus 2 for the “◦” matching a gap. So the alignment again yields a score of at least 10.

The statement about \( Y_{i,2} \) and \( Z \) can be proved in a similar fashion. \( \square \)

**Lemma 7.** Alignments of the pairs \( \{Y_{i,1}, X_{\alpha_{i,1}}\} \) and \( \{Y_{i,2}, X_{\alpha_{i,2}}\} \) yield scores of 6 and 3, respectively, if they fulfil property B. Violating property B yields scores of at least 8 and 5, respectively.

With the fact that an optimal alignment of a pair \( \{Y_{i,1}, Y_{i,2}\} \) has score 7 we can prove the following.

**Lemma 8.** Given an arbitrary alignment with score \( 31 \cdot t - 2 \cdot g \) we can construct a variable-consistent alignment with less or equal score in polynomial time.

*Proof.* Let \( \mathcal{A} \) be an arbitrary alignment with \( D_W(\mathcal{A}) = 31 \cdot t - 2 \cdot g \).

Let \( I \) be the set of all \( i \) such that \( Y_{i,1} \) and \( Y_{i,2} \) fulfil properties A and B. This implies an assignment for the variables \( U_I = \{x_j \in U | \exists i \in I : X_j \in S_i\} \). Let \( \overline{I} = \{1, \ldots, t\} \setminus I \). Because in every set \( S_i \) for \( i \in \overline{I} \) there exists a sequence \( Y_{i,q} \) that violates property A or B, we have \( D_i(\mathcal{A}) \geq 31 \) for each \( i \in \overline{I} \) due to Lemmas 6 and 7.

For \( i \in \overline{I} \) if \( x_{\alpha_{i,q}} \in U_I \) (\( q \in \{1,2\} \)), we realign \( Y_{i,q} \) with respect to \( x_{\alpha_{i,q}} \). Then we assign an arbitrary value to the variables in \( U \setminus U_I \) and realign the corresponding \( Y_{i,q} \) and \( X_j \).

By these modifications we obtain an alignment \( \mathcal{A}' \). Then \( D_i(\mathcal{A}') = D_i(\mathcal{A}) \) for \( i \in I \) and \( D_i(\mathcal{A}') \leq 31 \leq D_i(\mathcal{A}) \) otherwise. Thus, \( D_W(\mathcal{A}') \leq D_W(\mathcal{A}) \). \( \mathcal{A}' \) is variable-consistent due to its construction and can be computed in polynomial time. \( \square \)

The alignment obtained yields an assignment that satisfies at least \( g \) equations of \( \mathcal{G} \).

**Theorem 3.** BMSA is MAX SNP-hard.
Proof. We reduce MAX-E2-neg-Lin2 to BMSA. $f_1$ is given by the construction of $S$ from a family $G$ of $t$ equations. One can see that $\text{opt}(S) \leq 31 \cdot t$.

An equation of $G$ will be satisfied by 2 of the 4 possible assignments of its variables. Therefore, for every multiset $G$ of $t$ equations an assignment exists that satisfies at least $\frac{1}{2} \cdot t$ equations. Then for $\gamma_1 = 62$ we have $\text{opt}(S) \leq \gamma_1 \cdot \text{opt}(G)$.

Given an alignment of $S$ with score $31 \cdot t - 2 \cdot g'$ for some $g'$ we can find an assignment satisfying $g \geq g'$ equations of $G$ due to Lemma 8. Let $\gamma_2 = \frac{1}{2}$, then $|g - \text{opt}(G)| \leq \gamma_2 \cdot |(31 \cdot t - 2 \cdot g') - \text{opt}(S)|$ holds.

\[ \square \]

Theorem 4. BMSA has no polynomial time approximation with performance ratio $\frac{324}{323} - \epsilon$ for any $\epsilon > 0$, unless $\mathsf{NP} = \mathsf{P}$.

Proof. An instance of MAX-E2-neg-Lin2 consisting of $t$ equations is $\eta$-satisfiable iff the corresponding instance of BMSA has an alignment with score $(31 - 2 \cdot \eta) \cdot t$.

The optimal alignment of a BMSA instance corresponding to a $(\frac{18}{22} - \xi)$-satisfiable instance of MAX-E2-neg-Lin2 has score $(31 - 2 \cdot (\frac{18}{22} - \xi)) \cdot t = \frac{323 + 22 \cdot \xi}{11}$. Using the $(\frac{324}{323} - \epsilon)$-approximation algorithm for BMSA we are able to find an alignment with score at most $(\frac{324}{323} - \epsilon) \cdot \frac{323 + 22 \cdot \xi}{11} \cdot t =: K_1$.

The optimal alignment of a BMSA instance corresponding to a $(\frac{17}{22} + \xi)$-satisfiable instance of MAX-E2-neg-Lin2 has score $(31 - 2 \cdot (\frac{17}{22} + \xi)) \cdot t =: K_2$. We have $K_1 < K_2$ iff $\xi < \frac{1}{22} \cdot \frac{323^2 - \epsilon}{647 - 323 \cdot \epsilon}$. Choose $\xi$ with $0 < \xi < \frac{1}{22} \cdot \frac{323^2 - \epsilon}{647 - 323 \cdot \epsilon}$. Then the $(\frac{324}{323} - \epsilon)$-approximation for BMSA can be used to distinguish $(\frac{18}{22} - \xi)$- and $(\frac{17}{22} + \xi)$-satisfiable instances of MAX-E2-neg-Lin2. This would imply $\mathsf{NP} = \mathsf{P}$ due to Theorem 2.

\[ \square \]

Since WMSA is a generalization of BMSA it is also MAX $\mathcal{SNP}$-hard and we obtain the same non-approximability result.

6 Conclusions

We have shown MAX $\mathcal{SNP}$-hardness and proved a numerical lower bound for the approximability of weighted multiple sequence alignment (WMSA). These results hold even if we restrict the problem to binary weights (BMSA). Furthermore, BMSA and WMSA are equivalent with respect to their approximability. But the distance to the best known upper bound is huge. An obvious goal is to reduce this gap.

Finally, we would like to know how well the unweighted version of the multiple sequence alignment problem with metric SP-score can be approximated.

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A Greedy Algorithm for Optimal Recombination

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Abstract. Let $\Sigma$ be an alphabet and $\Sigma^n$ denote the collection of all sequences of length $n$ over $\Sigma$. For any $s_1 = a_1a_2 \cdots a_ia_{i+1} \cdots a_n$, $s_2 = b_1b_2 \cdots b_j b_{j+1} \cdots b_n \in \Sigma^n$, a recombination of $s_1$ and $s_2$ at position $j$ is defined as an operation that crosses $s_1$ and $s_2$ at position $j$ and generates $t_1 = a_1a_2 \cdots a_j b_{j+1} \cdots b_n$ and $t_2 = b_1b_2 \cdots b_j a_{j+1} \cdots a_n$. Denote $A$ and $S$ two collections of sequences. In this paper, we discuss generating $A$ from $S$ by a series of recombinations in minimum number of steps. We present a greedy algorithm for finding the optimal recombination evolutionary history from $S$ to any tree $A$ of sequences when $|S| = 2$.

1 Introduction

Various types of mutations on sequences play an important role in computational biology. Transformations using insertion, deletion, substitution, and reversal are widely studied by applying statistics and algorithms (cf. [1,2]). Recently, much attention has been paid to recombination of sequences. Hein developed some algorithms for recombination problems (cf. [3,4]). Kececioglu and Gusfield discussed a recombination distance problem on generating a third sequence from a given pair of sequences in optimal recombination cost (cf. [5]).

Generally, an edit distance problem of genomes is widely studied and its aim is to find the optimal evolutionary history from some ancestors to some descendents by using certain types of mutations such as insertion, deletion, substitution(point mutation),reversal, etc. Parsimony trees and phylogenetic trees are some interesting problems in the category. In [6], an alignment with recombination is discussed and it is an edit distance problem involved recombinations.

In this paper, we discuss a similar distance problem involved recombinations. The purpose is to generate a collection $A$ of sequences from another collection $S$ of sequences by a series of recombinations in minimum number of steps.

1.1 Problem and Example

Definition 1. Let $\Sigma$ be an alphabet (e.g., $\Sigma = \{a, c, g, t\}$, etc) and $\Sigma^n$ the collection of all sequences of length $n$ over $\Sigma$. For any $s_1 = a_1a_2 \cdots a_ia_{i+1} \cdots a_n$, $s_2 = b_1b_2 \cdots b_j b_{j+1} \cdots b_n \in \Sigma^n$, a recombination of $s_1$ and $s_2$ at position $j$ is defined as an operation that crosses $s_1$ and $s_2$ at position $j$ and generates $t_1 = a_1a_2 \cdots a_j b_{j+1} \cdots b_n$ and $t_2 = b_1b_2 \cdots b_j a_{j+1} \cdots a_n$. Denote $A$ and $S$ two collections of sequences. In this paper, we discuss generating $A$ from $S$ by a series of recombinations in minimum number of steps. We present a greedy algorithm for finding the optimal recombination evolutionary history from $S$ to any tree $A$ of sequences when $|S| = 2$.

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optimal recombination process generating $A$ is minimized among all such possible recombination processes.

**Definition 2.** Let $A, S \subseteq \Sigma^n$. $S$ is called a recombination ancestor of $A$ if $A$ can be generated from $S$ by a series of recombinations, i.e., there exist a series of recombinations $R_i(s_1^{(i)}, s_2^{(i)}, j^{(i)}, t_1^{(i)}, t_2^{(i)})$ ($i = 1, 2, \ldots, p$) such that (1) $s_1^{(i)}, s_2^{(i)} \in S \cup \{t_1^{(k)}, t_2^{(k)}| 1 \leq k \leq i - 1\}$, (2) $A \subseteq S \cup \{t_1^{(i)}, t_2^{(i)}| 1 \leq i \leq p\}$. $R_i(s_1^{(i)}, s_2^{(i)}, j^{(i)}, t_1^{(i)}, t_2^{(i)})$ $(i = 1, 2, \ldots, p)$ is called a recombination process generating $A$ from $S$. It is called optimal if the number $p$ of recombination steps is minimized among all such possible recombination processes.

**Recombination Problem 1.** Let $A \subseteq \Sigma^n$ and $S$ a recombination ancestor of $A$. Find the optimal recombination process generating $A$ from $S$. Denote $n(S, A)$ the number of optimal recombination steps generating $A$ from $S$.

**Example 1.** Let $A = \{a_1 = 01010, a_2 = 10001, a_3 = 10110, a_4 = 11001\}$ and $S = \{s_1 = 00101, s_2 = 11010\}$. We have $n(S, A) = 4$ and the following is an optimal recombination process generating $A$ from $S$.

$\begin{align*}
R_1(s_1, s_2; 1; a_1, b_4) & : \text{From } 00101 = b_1 = s_1 \\
& \quad \quad \underset{1 \to 1010}{\text{To } 0} \quad = b_2 = s_2 \\
R_2(b_4, s_2; 2; b_5, b_6) & : \text{From } 10101 = b_4 \\
& \quad \quad \quad \underset{10 \to 1010}{\text{To } 0} \quad = b_3 = a_1 \\
& \quad \quad \quad \quad \underset{1 \to 0101}{\text{To } 10} \quad = b_4 \\
R_3(b_4, s_2; 3; a_3, a_4) & : \text{From } 10101 = b_9 \\
& \quad \quad \quad \quad \underset{110 \to 10}{\text{To } 101} \quad = b_7 = a_3 \\
& \quad \quad \quad \quad \quad \underset{110 \to 01}{\text{To } 101} \quad = b_8 = a_4 \\
R_4(s_1, b_5; 3; a_2, b_9) & : \text{From } 00101 = b_1 \\
& \quad \quad \quad \quad \quad \underset{10 \to 10}{\text{To } 001} \quad = b_3 \\
& \quad \quad \quad \quad \quad \quad \underset{100 \to 10}{\text{To } 001} \quad = b_9 \\
& \quad \quad \quad \quad \quad \quad \quad \quad \underset{100 \to 01}{\text{To } 001} \quad = b_{10} = a_2
\end{align*}$

If $S$ consists of only two sequences, then Problem 1 can be regarded as one with $|\Sigma| = 2$. For binary case, $S = \{00 \ldots 0, 11 \ldots 1\}$ always generates an arbitrary binary sequence by recombinations. So it is a recombination ancestor for any
\[A \subseteq \{0, 1\}^n.\] Moreover, given any recombination ancestor \(S = \{s_{i1}s_{i2}\cdots s_{in}|i = 1, 2\}\) of \(A = \{a_{i1}a_{i2}\cdots a_{in}|i = 1, 2, \cdots, k\},\) for each \(1 \leq j \leq n,\) we define a function \(f_j\) on \(\{s_{1j}, s_{2j}\}\) \((s_{1j} \neq s_{2j})\) such that \(f_j(s_{1j}) = 0\) and \(f_j(s_{2j}) = 1.\) Then we get \(S_0 = \{f_1(s_{11})f_2(s_{12})\cdots f_n(s_{1n}), f_1(s_{21})f_2(s_{22})\cdots f_n(s_{2n})\} = \{00\cdots 0, 11\cdots 1\}\) and \(A_0 = \{f_1(a_{i1})f_2(a_{i2})\cdots f_n(a_{in})|i = 1, 2, \cdots, k\}.\) And \(n(S, A) = n(S_0, A_0).\) For \(|\Sigma| = 2,\) Recombination Problem 1 is reduced to the following simple form.

**Recombination Problem 2.** Let \(A \subseteq \{0, 1\}^n\) be an arbitrary collection of binary sequences and \(S = \{00\cdots 0, 11\cdots 1\}.\) Find the optimal recombination process generating \(A\) from \(S.\)

In this paper, we discuss Problem 2 and find algorithms for it.

### 1.2 Terminology and Notation

**Definition 3.** Let \(a = a_1a_2\cdots a_p\cdots a_q\cdots a_n \in \{0, 1\}^n.\) \(I = a[p, q] = a_pa_{p+1}\cdots a_q\) is called an alternative segment of \(a\) if \(a_i \neq a_{i+1}\) for all \(p \leq i \leq q - 1.\) An alternative block is a maximal alternative segment. Denote \(I_a = \{I_1, I_2, \cdots, I_k\}\) the collection of all alternative blocks of \(a.\) Define the core of \(a\) as \(C_a = a[s, t],\) which is the minimum segment of \(a\) containing all \(I_i(1 \leq i \leq k).\) Define the length of \(a[p, q]\) as \(l(a[p, q]) = q - p\) and the length of \(I_a\) as \(l(I_a) = \sum_{j=1}^{k} l(I_j).\) Denote \(P(A) = \{p|1 \leq p \leq n - 1, a_p \neq a_{p+1}\}\) for some \(a \in A.\) For example, if \(a = 11101011110000100,\) then \(I_1 = a[3, 7] = 10101, I_2 = 1010, I_3 = 010.\) \(I_a = \{I_1, I_2, I_3\}.\) \(C_a = a[3, 17] = 10101111000010.\) \(l(I_1) = 4\) and \(l(I_a) = 9.\)

**Definition 4.** Let \(a, b \in \{0, 1\}^n\) with \(C_a = a[p, q]\) and \(C_b = b[u, v].\) If \(u \leq p \leq q \leq v\) and \(a[p, q] = b[p, q],\) we say that \(b\) covers \(a,\) denoted by \(b \subseteq a.\) \(b\) and \(a\) are called disjoint or independent if \(q \leq u,\) or \(v \leq p,\) denoted by \(a \cap b = \emptyset.\) If \(q \leq u,\) we say \(a < b.\) For any alternative segments \(I\) and \(J,\) we similarly define \(I \subseteq J, I \cap J = \emptyset,\) and \(I < J.\)

**Definition 5.** Denote \(C(A) = \{C_a|a \in A\}.\) \(A\) is called a nest of sequences if \(a \subseteq b,\) or \(b \subseteq a\) for any \(a, b \in A.\) It is called a tree of sequences if either \(a \cap b = \emptyset,\) or \(a \subseteq b,\) or \(b \subseteq a\) for any \(a, b \in A.\) For \(a, b \in A,\) if \(b \subseteq a\) and there exists no \(c \in A\) such that \(b \subseteq c \subseteq a\) with \(c \neq a\) and \(c \neq b,\) then \(b\) is called a branch of \(a.\) Denote \(B_a\) the collection of all branches of \(a\) in \(A.\) A branch is called a leaf \((\text{or longest branch})\) if it has no branch in \(A.\) \(a \in A\) is called a root \((\text{or oldest branch})\) if it is not a branch of any other sequences in \(A.\)

### 1.3 Related Work

Our recombination problem is a generalization of the following problem (cf. [5]).

**Recombination Distance.** Given \(a = a_1a_2\cdots a_m,\) \(b = b_1b_2\cdots b_n,\) and \(c = c_1c_2\cdots c_k,\) find the minimum cost recombination to produce \(c\) from \(a\) and \(b.\)
The goal is to generate a third sequence from a given pair by recombinations consisting of multiple crossovers and point mutations. In our problems, we define a recombination consisting of a single crossover with no point mutations. We discuss a more general problem to construct the optimal recombination spanning history from one family of sequences to another one. Its general case with multiple crossovers and point mutations is NP-complete (cf. [5, 6]).

2 Theorems and Algorithms

In this section, we show some theorems on optimal recombination processes and design a greedy algorithm for finding the optimal recombination process for a tree of binary sequences. We always assume $S = \{00 \cdots 0, 11 \cdots 1\}$.

**Theorem 1.** Let $a \in \{0, 1\}^n$. Then $n(S, \{a\}) = l(I_a)$.

**Proof.** $a$ is optimally generated by $R(s_1, s_2; p; b_1, c_1), R(b_1, c_1; p + 1; b_2, c_2), R(b_2, c_2; p + 2; b_3, c_3), \cdots$, $R(b_q-p+1, c_q-p-1; q-1; b_{q-p}, c_{q-p}), \cdots$ over all positions in $P(\{a\})$. Therefore, $n(S, \{a\}) = l(I_a)$.

The series of recombinations in the proof is called a recombination extension and denoted by $\text{ext}(s_1, a)$. Generally, if $a \subseteq b$ and $a$ has been generated, then $b$ can be generated by series of recombinations on the positions in $P(\{b\})$ but not $P(\{a\})$, called an recombination extension from $a$ to $b$ and denoted by $\text{ext}(a, b)$.

**Theorem 2.** If $A$ is a nest, then $n(S, A) = |P(A)|$.

**Proof.** By Theorem 1 and recombination extensions.

**Theorem 3.** Let $A_1, A_2 \subseteq \{0, 1\}^n$ be independent nests, i.e., $a \cap b = \emptyset$ for any $a \in A_1, b \in A_2$. Then $n(S, A_1 \cup A_2) = |P(A_1)| + |P(A_2)|$.

**Proof.** Apply Theorem 2 to $A_1$ and $A_2$, respectively.

**Theorem 4.** Let $A_1, A_2 \subseteq \{0, 1\}^n$ be independent nests and $a \in \{0, 1\}^n$. For any $b_1 \in A_1$ and $b_2 \in A_2$, (1)$b_1 < b_2$ and (2)$b_1, b_2 \subseteq a$. Then $n(S, A_1 \cup A_2 \cup \{a\}) = |P(\{a\})| + 1$.

**Proof.** Choose a position $p$ between $A_1$ and $A_2$, then $C_a$ is partitioned into $C_1$ and $C_2$. Choose $a_1$ and $a_2$ with $C_{a_1} = C_1$ and $C_{a_2} = C_2$. Then the theorem follows by applying Theorem 2 to both $A_1 + \{a_1\}$ and $A_2 + \{a_2\}$ with one more recombination $R(a_1, a_2; p)$, i.e., $|P(\{a_1\})| + |P(\{a_2\})| + 1 = |P(\{a\})| + 1$ steps.

**Theorem 5.** Let $A$ be a tree. Then $n(S, A) = |P(A)| + \sum_{a \in A}(|B_a| - 1)$.

**Proof.** By Theorem 4 and induction on $|A|$. Choose a root $a \in A$ and denote $B_a = \{d_1 < d_2 < \cdots < d_p\}$. We define $\text{ext}(d_1, d_2, \cdots, d_p; a)$ as the following recombination process: (1) Choose a position $p_i$ between $d_i$ and $d_{i+1}$ $(1 \leq i \leq p - 1)$, (2) Partition $C_a$ by all $p_i$ into $C_i$ and choose $b_i$ with $C_{b_i} = C_i(1 \leq i \leq p)$, (3) Make $\text{ext}(d_i, b_i)$ to get $b_i$ from $d_i(1 \leq i \leq p)$, (4) Make the recombinations: $R(b_{i-1}, b_i; p_i; f_i, g_i) (2 \leq i \leq p - 1)$ with $b_j \subseteq f_i (1 \leq j \leq i - 1)$. Then $f_{p-1} = a$. $\text{ext}(d_1, d_2, \cdots, d_p; a)$ generates $a$ from $d_1, d_2, \cdots, d_p$ in optimal steps. $A$ is then generated from branches to roots by induction.
Based on Theorem 5, we now design a greedy algorithm for finding the optimal recombination spanning evolutionary history generating a tree $\mathcal{A}$ from $\mathcal{S}$.

**Greedy Algorithm**  
**Input:** $\mathcal{A}$ (nest/tree).  
**Output:** Optimal recombination process.  
Step 1 Find $\mathcal{B}_a$ for all $a \in \mathcal{A}$ and set $E = \emptyset$.  
Step 2 While $\mathcal{A} - E \neq \emptyset$ {  
Choose a leaf (youngest branch) $a \in \mathcal{A} - E$ and denote $\mathcal{B}_a = \{d_1 < d_2 < \cdots < d_p\}$.  
Generate $a$ by $\text{ext}(d_1, d_2, \cdots, d_p; a)$.  
$E = E + \{a\}$.  
}

**Theorem 6.** Greedy Algorithm finds the optimal recombination process for a tree $\mathcal{A}$ and the run time is $O(|\mathcal{A}|^2n)$.

**Proof.** For each pair $a$ and $b$ of sequences, the algorithm compares $n$ positions to determine whether $a \subseteq b$ or not. There are $O(|\mathcal{A}|^2)$ pairs of sequences. Therefore the run time is $O(|\mathcal{A}|^2n)$.

### 3 Conclusion

The greedy algorithm is designed to generate a tree from two sequences. The most interesting part of the problem is to generate an arbitrary collection of sequences from a given recombination ancestor. Our discussion may be applied to some problems in human SNP (single nucleotide polymorphism) genome project.

### References

Generating Well-Shaped $d$-dimensional Delaunay Meshes

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Abstract. A $d$-dimensional simplicial mesh is a Delaunay triangulation if the circumsphere of each of its simplices does not contain any vertices inside. A mesh is well-shaped if the maximum aspect ratio of all its simplices is bounded from above by a constant. It is a long-term open problem to generate well-shaped $d$-dimensional Delaunay meshes for a given polyhedral domain. In this paper, we present a refinement-based method that generates well-shaped $d$-dimensional Delaunay meshes for any PLC domain with no small input angles. Furthermore, we show that the generated well-shaped mesh has $O(n^d)$ simplices, where $n$ is the smallest number of $d$-simplices of any almost-good meshes for the same domain. A mesh is almost-good if each of its simplices has a bounded circumradius to the shortest edge length ratio.

Keywords: Mesh generation, Delaunay triangulation, well-shaped, aspect ratio, radius-edge ratio, computational geometry, algorithms.

1 Introduction

Mesh generation is the process of breaking a geometry domain into a collection of primitive elements. In this paper we exclusively consider $d$-dimensional simplicial Delaunay meshes. The aspect-ratio of a mesh is the maximum aspect-ratio among all of its simplicial elements. A mesh is well-shaped if its aspect ratio is bounded from above by a small constant. The aspect ratio of a simplex is usually defined as the ratio of its circumradius to its inradius. Generating well-shaped $d$-dimensional Delaunay meshes is one of the long-term open problems in mesh generation when $d > 2$. An alternative but weaker quality measurement is to use the radius-edge ratio. It is the circumradius divided by the shortest edge length of the simplex. The radius-edge ratio of a mesh is the maximum radius-edge ratio among all of its elements. A mesh is almost-good if it has a small radius-edge ratio. Numerous methods guarantee to generate almost-good 3-dimensional Delaunay meshes.

Bern et al. showed that any set of $n$ $d$-dimensional points had a Steiner Delaunay triangulation with $O(n^{\lceil d/2 \rceil})$ simplices, none of which has an obtuse dihedral angle. No bound depending only on $n$ is possible if we require a bound on the minimum dihedral angle. We assume a $d$-dimensional piecewise linear complex (PLC) domain $\Omega$ as the input. Shewchuk showed that if each
$k$-dimensional constraining facet in $\Omega$ with $k \leq d - 2$ is a union of strongly Delaunay $k$-simplices, then $\Omega$ has a $d$-dimensional constrained Delaunay triangulation. The Delaunay refinement method \cite{3,9,10} can be extended to $d$-dimensions to generate almost-good Delaunay meshes if $\Omega$ satisfying the projection lemma \cite{10} and having small input angles. However, it encounters significant difficulties in producing well-shaped Delaunay meshes in 3D. Chew \cite{4} first proposed to add a vertex inside the circumsphere of any given badly-shaped tetrahedron (sliver) to remove it. The difficulty is in proving the existence of, and finding, a Steiner vertex that does not itself result in slivers. To make it possible, Chew defined a tetrahedron to be a sliver if it has a large aspect ratio, a small radius-edge ratio and its circumradius is no more than one unit. Without the restriction on the circumradius length, almost all results in \cite{4} donot hold any more. In addition, with this restriction, the termination of his algorithm is straightforward. Recently, Li \cite{6} extended this algorithm and showed how to generate well-shaped non-uniform Delaunay meshes in 3D.

In this paper, we show how to generate $d$-dimensional well-shaped meshes for a PLC domain with no small angles. In Section 2, we review the basic concepts and define what is a $d$-dimensional sliver simplex. Then we give an algorithm in Section 3 to generate $d$-dimensional well-shaped Delaunay meshes. It basically adds a point around the circumcenter of each of $d$-simplices containing any $k$-dimensional sliver (hereafter, $k$-sliver). We prove its correctness in Section 4 and its termination guarantee in Section 5 by showing that the distance between the closest mesh vertices is just decreased by a constant factor compared with that of the input mesh. In Section 6, we show that the size of the generated mesh is within a constant factor of the size of any almost-good mesh generated for the same domain. Section 7 concludes the paper with discussions.

2 Preliminaries

After inserting a new vertex $p$, every new $d$-dimensional simplex created in the Delaunay triangulation of the new vertex set has $p$ as one of its vertices. The new triangulation can be updated by efficient operations local to the vertex $p$. A sphere centered at a point $c$ with a radius $r$ is denoted as $(c, r)$ hereafter. It is called empty if it does not contain any mesh vertices inside. The nearest neighbor graph defined by a $d$-dimensional vertex set is contained in the Delaunay triangulation of the vertex set. Thus the shortest edge length of the Delaunay triangulation is the closest distance among mesh vertices. This fact is used in proving the termination guarantee of our algorithm.

Delaunay refinement methods have been shown to be effective in generating almost-good meshes in 2 and 3 dimensions \cite{9,10}. There is also no much difficulty to extend them to $d$-dimensions, if the input domain satisfies a projection lemma \cite{10}, to generate meshes with radius-edge ratio no more than $\sqrt{2}^{d-1}$.

\footnote{Surprisingly, a simple extension of 3-dimensional sliver definition to $d$-dimensions does not work here.}
For a \( k \)-simplex \( \mu \), its min-circumsphere is the smallest \( d \)-dimensional sphere containing the vertices of \( \mu \) on its surface. A point encroaches the domain boundary if it is contained inside the min-circumsphere of a boundary \( k \)-simplex \( \mu \). Here a \( k \)-simplex \( \mu \) is a boundary simplex if it belongs to the Delaunay triangulation of a \( k \)-dimensional input boundary polyhedral face. A \( k_1 \)-simplex \( \tau_1 \) directly encroaches another \( k_2 \)-simplex \( \tau_2 \) if the circumcenter of \( \tau_1 \) encroaches \( \tau_2 \). A \( k_1 \)-simplex \( \tau_1 \) indirectly encroaches another \( k_2 \)-simplex \( \tau_2 \) if the circumcenter of \( \tau_1 \) encroaches a simplex \( \mu \) and \( \mu \) directly or indirectly encroaches \( \tau_2 \). Assume the circumcenter of a \( k_1 \)-simplex \( \tau \) encroaches a boundary \( k_2 \)-simplex \( \mu \). Call \( \tau \) the encroaching simplex and \( \mu \) the encroached simplex. Assume that \( \mu \) contains the projection of the circumcenter of \( \tau \) inside. Then Shewchuk [10] showed that the circumradii of \( \tau \) and \( \mu \) satisfy that \( R_\mu \geq \frac{1}{\sqrt{2}} R_\tau \).

A \( d \)-simplex is bad if it has a large aspect ratio. Let’s consider a \( k \)-simplex \( \tau \), where \( 1 \leq k \leq d \). For later convenience, we use \( R_\tau, L_\tau \) and \( \rho(\tau) = R_\tau / L_\tau \) to denote the circumradius, the shortest edge length and the radius-edge ratio of \( \tau \). Let \( V \) be its volume. We define \( \sigma = \sigma(\tau) = V / L_\tau^k \) as a measure of its quality. Let \( \varrho_0 \) and \( \sigma_k \), \( 1 \leq k \leq d \) be positive constants that we specify later.

**Definition 1.** A \( k \)-simplex \( \tau \) is well-shaped if \( \rho(\tau) \leq \varrho_0 \) and \( \sigma(\tau) \geq \sigma_k \).

**Definition 2.** [Sliver] Call a \( k \)-simplex \( \tau \) a sliver if \( \rho(\tau) \leq \varrho_0 \), \( \sigma(\tau) < \sigma_k \), and each of its facets is well-shaped if \( k > 3 \).

We call a \( d \)-simplex \( \tau \) a sliver-simplex if it contains a \( k \)-sliver. If a \( d \)-simplex \( \tau \) has a small radius-edge ratio and a small \( \sigma \) value, then it must have a face \( \chi \) that is a sliver or \( \tau \) itself is a \( d \)-sliver. It is not difficulty to prove that the volume of a \( d \)-simplex \( \tau \) is at most \( \varphi_d R_\tau^d \), where \( \varphi_d = 2 \prod_{i=2}^d \frac{(i-1)(i-1)!2^{(i+1)}/(i+1)!}{i^{i+1}} \). The aspect ratio of \( \tau \) is then at most \( \frac{(1+d)\varphi_d \rho(\tau)^d}{\sigma(\tau)} \). This verifies our definition of \( \sigma \).

### 3 Refinement Algorithm

For a \( k \)-simplex \( \tau \), we call the intersection of its min-circumsphere with the affine space defined by its vertices as its \( k \)-sphere. When refine a \( k \)-simplex \( \tau \), we add a point \( p \) inside the shrinked \( k \)-sphere \( (c_\tau, \delta R_\tau) \), where \( \delta < 1 \) is a constant. The point \( p \) is good if its insertion will not introduce any small slivers in the new Delaunay triangulation. Here a created sliver \( \mu \) is small if \( R_\mu \leq b R_\tau \) for a constant \( b \) specified later. We call the solid \( k \)-dimensional ball \( (c_\tau, \delta R_\tau) \) the picking region of \( \tau \), denoted by \( P(\tau) \). Its volume is \( \phi_k \delta^k R_\tau^k \), where \( \phi_k = \frac{2\pi^{k/2}}{k!} \).

**Algorithm:** REFINE \((\varrho_0, \sigma_0, \delta, b)\)

**Enforce Empty Encroachment:** Add the circumcenter \( c_\tau \) of any encroached boundary simplex \( \tau \) and update the Delaunay triangulation. If \( c_\tau \) encroaches any lower dimensional boundary simplex \( \mu \), add \( c_\mu \) instead of \( c_\tau \).

**Clean Large Radius-Edge Ratio:** Add the circumcenter \( c_\tau \) of any \( d \)-simplex \( \tau \) with a large \( \rho(\tau) \) and update the Delaunay triangulation. If \( c_\tau \) encroaches any boundary \( k \)-simplex, we apply the last rule instead of adding \( c_\tau \).
Fig. 1. Left: a 4-dimensional sliver example; Middle: the picking region of a simplex; Right: the forbidden region of $\mu$ with circumradius $Y$. Here $R = 2\rho_0 Y$ and $D = 2k\sigma_0 Y$.

**Clean the Slivers:** For a sliver-simplex $\tau$, add a good point $p \in \mathcal{P}(\tau)$ and update the Delaunay triangulation. If the circumcenter $c_\tau$ encroaches the domain boundary, we apply the following rules instead of finding $p$. If the insertion of $p$ introduces some new $d$-simplices with large radius-edge ratio, we apply the previous rule to eliminate them immediately.

**Encroach Boundary Simplices:** If a boundary $k$-simplex $\mu$ is encroached directly or indirectly by a $d$-simplex with a large radius-edge ratio, add the circumcenter $c_\mu$ and update the Delaunay triangulation. However, if $c_\mu$ encroaches any other lower dimensional boundary simplex $\mu_1$, we insert the circumcenter of $\mu_1$ instead of adding $c_\mu$.

If a boundary $k$-simplex $\mu$ is encroached directly or indirectly by a sliver-simplex, add a good point $p \in \mathcal{P}(\mu)$ and update the Delaunay triangulation. However, if $c_\mu$ encroaches any other lower dimensional boundary simplex $\mu_1$, we add a good point from $\mathcal{P}(\mu_1)$ instead of adding $p$.

The key part of the algorithm is to find a good point $p$ to refine a sliver-simplex $\tau$. We select some $k$ random points from $\mathcal{P}(\tau)$. Then choose the point that optimizes the quality $\sigma$ of all created small simplices with a radius-edge ratio less than $\rho_0$. Another approach is based on a randomized selection [4,6,7]. We randomly select a point $p \in \mathcal{P}(\tau)$ until we find a good point $p$. By defining sliver and small slivers properly, we can show that we are expected to find a good point $p$ in constant rounds. Then the rest of the paper is devoted to prove the termination guarantee and the good grading guarantee of the algorithm.

## 4 Proofs of Correctness

### 4.1 Sliver Regions

Recall that a $k$-simplex $\tau$ is a sliver if $\rho(\tau) \leq \rho_0$, $\sigma(\tau) \leq \sigma_k$, and all of its facets are well-shaped. The quality measure $\sigma(\tau)$ is related with a distance-radius ratio defined in [25]. Consider any vertex $p$ of $\tau$. Let $\mu$ be the facet formed by other vertices of $\tau$. Let $D$ be the Euclidean distance of point $p$ from the hyperplane passing through $\mu$. Recall $R_\mu$ is the circumradius of $\mu$. 


Lemma 1. For any k-simplex \( \tau \), \( \frac{D}{R_\mu} \leq 2k\frac{\sigma(\tau)}{\sigma(\mu)} \).

**Proof.** It follows from the volume of \( \tau \): \( V = \sigma(\tau)L_\tau^k = \frac{D}{R_\mu} \sigma(\mu)L_\mu^{k-1} \).

A k-simplex \( \tau \) is a sliver if \( \rho(\tau) \leq \varrho_0 \) and \( \sigma(\tau) = \frac{V}{L_\tau^k} \leq \sigma_0^k \) for a small constant \( \sigma_0 \), i.e., \( \sigma_k = \sigma_0^k \). Consequently, we have \( \frac{D}{R_\mu} \leq 2k\sigma_0 \) if \( \tau \) is a k-sliver.

Given a well-shaped k-simplex \( \tau \), a point \( p \) that forms a \((k+1)\)-sliver together with \( \tau \) can not be anywhere. It must be around the circumsphere of \( \tau \) and is not far-away from the hyperplane passing \( \tau \). We call the locus of such point \( p \) as the sliver region or forbidden region \( F_\tau \) of the simplex \( \tau \). It is easy to show that \( F_\tau \) is contained in the solid sun-hour glass shaped region as illustrated in the right figure of Figure 1. Let \( |F_\tau| \) denote the volume of \( F_\tau \).

**Lemma 2.** For any well-shaped k-simplex \( \tau \), \( |F_\tau| \leq c_d\sigma_0^{d-k+1}R_\tau^d \), where \( c_d \) is a constant depending only on \( \varrho_0, k, \) and \( d \).

**Proof.** We know that \( F_\tau \) is inside the slab region defined by the two hyperplanes \( H_1 \) and \( H_2 \) illustrated by the right figure of Figure 1. Let \( r_2, r_1 \) be the circumradius of the two spheres intersected by \( H_1 \) and the circumsphere of \( \tau \). Here \( r_2 = ||cp|| \) and \( r_1 = ||cq|| \). It is easy to show that \( r_2^2 \leq (1 - (2k\sigma_0)^2 + 8k\sigma_0\varrho_0) \cdot Y^2 \) and \( r_2^2 \geq (1 - (2k\sigma_0)^2 - 8k\sigma_0\varrho_0) \cdot Y^2 \). The volume of a k-sphere with the radius \( r \) is \( \phi_k r^k \). Then

\[
|F_\tau| \leq \phi_k(r_2^k - r_1^k) \cdot \phi_{d-k}D^{d-k}.
\]

Notice \( \forall i \geq 1, r_{i+2}^2 - r_{i+1}^2 \leq (r_i^2 + r_i^2)(r_{i+1}^2 - r_i^2) \). In addition, \( r_2^2 + r_1^2 = 2(1 - (2k\sigma_0)^2)Y^2 \leq 2Y^2 \), and \( r_2^2 - r_1^2 \leq 16k\varrho_0\sigma_0Y^2 \). The fact that \( r_2 \geq Y \) implies that \( r_2 - r_1 \leq 16k\varrho_0\sigma_0Y \). By induction, we have \( r_{i+1}^2 - r_i^2 \leq 2^{[k/2]+3}k\varrho_0\sigma_0Y^k \). Consequently, we have \( |F_\tau| \leq \phi_k \phi_{d-k}2^{[k/2]+3}k\varrho_0\sigma_0(2k\sigma_0)^d-kY^d \). The lemma follows by setting \( c_d = \phi_k \phi_{d-k}2^{[k/2]+3}k\varrho_0(2k)^{d-k} \).

A good point \( p \) used to refine a k-simplex \( \tau \) cannot be inside \( F_\mu \cap \mathcal{P}(\tau) \) for any well-shaped m-simplex \( \mu \). We use \( F_{\mu,k} \) to denote such intersection region.

**Lemma 3.** For any k-simplex \( \tau \) and a well-shaped m-simplex \( \mu \), \( |F_{\mu,k}| \leq c_k \sigma_0 R_\mu^k \), where \( c_k \) is a constant that depends only on \( \varrho_0, d \) and \( k \).

### 4.2 Existence

We then show that given a k-simplex \( \tau \) in an almost-good d-dimensional simplicial Delaunay mesh, there is a good point \( p \) in \( \mathcal{P}(\tau) \). Let \( \mathcal{S}(\tau) \) be the set of simplices in any dimensions that each can form a small Delaunay sliver with a point from \( \mathcal{P}(\tau) \). A created sliver \( \mu \) is Delaunay if it belongs to the Delaunay triangulation by inserting a point from \( \mathcal{P}(\tau) \); it is small if \( R_\mu \leq bR_\tau \).

We first recall some results from Talmor et al. and extend it to d-dimensions. For a mesh vertex \( v \), the **edge length variation**, denoted by \( \nu(v) \), is defined as the length of the longest edge incident on it divided by the length of the shortest edge incident on it. Talmor[12] proved that, given an almost-good mesh in 3D, the edge length variation of each mesh vertex is at most a constant depending on the radius-edge ratio of the mesh. This result is extended to d-dimensions.
Lemma 4. Given any vertex $v$ of a $d$-dimensional almost good mesh, $\nu(v) \leq \nu_0$, where $\nu_0$ is a constant depending on the mesh’s radius-edge ratio $\varrho_0$.

Lemma 5. Given a $k$-simplex $\tau$ of an almost-good mesh, the number of the vertices of $S(\tau)$ is at most a constant depending on $\nu_0$ and $b$.

Proof. Assume that a well-shaped $m$-simplex $\mu$ forms a Delaunay sliver together with a point $p \in \mathcal{P}(\tau)$. The edges of $\mu$ have lengths at most $2R_\mu \leq 2bR_\tau$. The edges incident on any vertex $q$ of $\mu$ before point $p$ is introduced have length at least $2bR_\tau/\nu_0$. It implies that the closest distance among all vertices from $S(\tau)$ is at least $2bR_\tau/\nu_0$. It is simple to show that all such vertices are inside the sphere centered at $c_\tau$ with a radius $(\delta + 2b)R_\tau$. Then by a volume argument, the number of vertices of $S(\tau)$ is a constant. 

Thus the number of simplices in $S(\tau)$ is also a constant, let’s say, $W$.

Theorem 1. Good point exists in any $k$-simplex $\tau$ of an almost-good mesh.

Proof. Each simplex $\mu$ from $S(\tau)$ claims a forbidden region $F_\mu$ with volume at most $c_k \cdot \sigma_0 R_\mu^k$. The circumradius $R_\mu$ of $\mu$ is at most $bR_\tau$. The volume of $\mathcal{P}(\tau)$ is $\phi_k \delta^k R_\tau^k$. Therefore, if we select $\sigma_0$ such that $W \cdot c_k \sigma_0 b^k R_\tau^k < \phi_k \delta^k R_\tau^k$, then the pigeonhole principal will guarantee the existence of a good point $p \in \mathcal{P}(\tau)$.

5 Termination Guarantee

A $k$-sliver is called created if it contains at least one Steiner vertex; otherwise it is called original. We classify the bad $d$-simplices into three categories. The simplices containing an original $k$-sliver are called original sliver-simplices. The simplices containing only created $k$-slivers are called created sliver-simplices. The third category has $d$-simplices with a large radius-edge ratio. Let $e(\tau)$ be the shortest edge length introduced by eliminating a $d$-simplex $\tau$. Here $e(\tau)$ could be less than $L_\tau$ for a sliver-simplex $\tau$. For simplicity, we assume that all original sliver-simplices are removed first.

Lemma 6. For an original sliver-simplex $\mu$, $e(\mu) \geq \frac{(1-\delta)}{\sqrt{2d+1}} L_\tau$.

Lemma 7. For a simplex $\tau$ with $\frac{R_\tau}{L_\tau} > \varrho_0$, $e(\tau) \geq \frac{\varrho_0}{\sqrt{2d-1}} L_\tau$.

The proofs are similar to the 3-dimensional counterpart, which are omitted. It remains to show that refining any created sliver-simplex will not introduce shorter edges to the mesh. Let’s consider a sliver-simplex $\mu$. Assume it contains a $k$-sliver $\tau$ created by inserting a point from the picking region of an element, say $f(\tau)$. Element $f(\tau)$ is called the parent of the $k$-sliver $\tau$. There are three cases: $f(\tau)$ is a sliver-simplex; $f(\tau)$ is a $d$-simplex with a large radius-edge ratio; $f(\tau)$ is an encroached boundary simplex.
Lemma 8. Assume a sliver-simplex \( \mu \) contains a \( k \)-sliver \( \tau \) created by splitting another sliver-simplex \( f(\tau) \). Then \( e(\mu) \geq \frac{(1-\delta)b}{\sqrt{2}^{d+1}} \cdot L_f(\tau) \).

Lemma 9. Assume a sliver-simplex \( \mu \) contains a \( k \)-sliver \( \tau \), and parent \( f(\tau) \) has a large radius-edge ratio. Then \( e(\mu) \geq \frac{(1-\delta)g_0}{\sqrt{2}^{d+1}} \cdot L_f(\tau) \).

Lemma 10. Assume a sliver-simplex \( \mu \) contains a sliver \( \tau \) created by splitting a boundary \( m \)-simplex \( \chi \) that is encroached directly or indirectly by \( p(\chi) \). Then

1. \( e(\mu) \geq \frac{(1-\delta)g_0L_{p(\chi)}}{2^{d+1}} \), if \( d \)-simplex \( p(\chi) \) has a large radius-edge ratio.
2. \( e(\mu) \geq \frac{(1-\delta)bL_{\chi}}{\sqrt{2}^{d+1}} \), if \( d \)-simplex \( p(\chi) \) is a sliver-simplex.

Combining all above analysis, we know that the shortest distance between all mesh vertices is at least \( \frac{1-\delta}{\sqrt{2}^{d+1}} \) factor of that of the original mesh if we choose \( b, \delta \) and \( g_0 \) such that \( (1-\delta)b \geq \sqrt{2}^{d+1} \) and \( (1-\delta)g_0 \geq 2^d \). We then have the following theorem by a volume argument.

**Theorem 2.** The algorithm terminates in generating well-shaped Delaunay meshes.

### 6 Good Grading Guarantee

As \([6,7,10]\), we study the relation between the nearest neighbor function defined by the final mesh and the local feature size function defined by the input domain. The local feature size \( lfs(x) \) of a point \( x \) is the radius of the smallest ball intersecting two non-incident segments or vertices. The nearest neighbor \( N(x) \) of a point \( x \) is its distance to the second nearest mesh vertex.

With each vertex \( v \), we associate an insertion edge length \( e_v \) equal to the length of the shortest edge connected to \( v \) immediately after \( v \) is introduced into the Delaunay mesh. Notice that \( v \) may not have to be inserted into the mesh actually. If \( v \) is an input vertex, then \( e_v \) is the distance between \( v \) and its nearest input neighbor. So \( e_v \geq lfs(v) \) from the definition of \( lfs(v) \). If \( v \) is from the picking region of a simplex \( \mu \), if simplex \( \mu \) has \( \rho(\mu) \geq g_0 \), then \( p(v) \) is the most recently inserted endpoint of the shortest edge of \( \mu \). If \( \mu \) is a sliver-simplex containing an original \( k \)-sliver \( \tau \), then \( p(v) \) is an end point of the shortest edge of \( \tau \). If \( \mu \) is a created sliver-simplex containing a created \( k \)-sliver \( \tau \), then \( p(v) \) is the most recently inserted vertex of \( \tau \). If \( \mu \) is an encroached boundary \( k \)-simplex, then \( p(v) \) is the encroaching vertex. For the sake of simplicity, we always use the almost-good
mesh generated by Delaunay refinement method as input mesh. Therefore the
boundary simplices can not be encroached by input vertices. Ruppert \cite{9} and
Shewchuk \cite{10} showed that $N()$ defined on the mesh generated by Delaunay
refinement method is within a constant factor of $lfs(v)$, i.e., $N(v) \sim lfs(v)$.

\textbf{Lemma 11.} Let $v$ be a vertex in the final mesh and let $p = p(v)$. Then we have
e$_v \geq lfs(v)$ for an input vertex $v$; and $e_v \geq C \cdot e_p$ for a Steiner point $v$, where $C$
is a constant specified in the proof.

\textbf{Proof.} If $v$ is an original input vertex, then $e_v \geq lfs(v)$ from the definition
of $lfs(v)$. Thus the theorem holds. Then consider a non-input vertex $v$. We first
consider that $v$ is selected from the picking region of a $d$-simplex; say $\mu$.

Case 1.1: $\mu$ is a $d$-simplex with large radius-edge ratio $\rho(\mu) \geq g_0$. The parent
$p$ is one of the end points of the shortest edge of $\mu$. Let $L_\mu$ be the length of
the shortest edge $pq$ of $\mu$. Then $q$ is an original vertex or is inserted before $p$.
Therefore, $e_p \leq \|p - q\| = L_\mu \leq \frac{L_\mu}{g_0}$. Thus $e_v = R_\mu \geq g_0 \cdot e_p$.

Case 1.2: $\mu$ is a sliver-simplex containing an original $k$-simplex $\tau$. The parent
$p = p(v)$ is one of the end points of the shortest edge of $\tau$. Let $L_\tau$ be the length
of the shortest edge $pq$ of $\tau$. Similar to the previous case, we have $e_p \leq \|p - q\| =
L_\tau$. Notice that $R_\mu \geq R_\tau \geq L_\tau/2$. Thus $e_v \geq (1 - \delta)R_\mu \geq \frac{1 - \delta}{2} \cdot e_p$.

Case 1.3: $\mu$ is a sliver-simplex containing a created $k$-simplex $\tau$. There are three
cases about the parent element $f(\tau)$ of $\tau$: $f(\tau)$ is a sliver; $f(\tau)$ has a large radius-
edge ratio; $f(\tau)$ is an encroached $k$-dimensional boundary simplex. Recall that the
parent vertex $p = p(v)$ is the most recently inserted vertex of $\tau$.

Subcase 1.3.1: $f(\tau)$ is a sliver. Recall that the insertion of $p$ from the picking
region of sliver $f(\tau)$ will always avoid creating small slivers. Thus, $R_\tau \geq bR_{f(\tau)}$;
where $R_{f(\tau)}$ is the circumradius of $f(\tau)$. Notice that $e_p \leq (1 + \delta)R_{f(\tau)}$. Thus,
e$_v \geq (1 - \delta)bR_{f(\tau)} \geq \frac{1 - \delta}{1 + \delta}b \cdot e_p$.

Subcase 1.3.2: $f(\tau)$ has $\rho(f(\tau)) \geq g_0$. Let $pq$ be an edge of simplex $\tau$, where
$p$ is inserted from the picking region of $f(\tau)$. Notice that $e_p \leq \|p - q\|$. We also
have $R_\tau \geq \|p - q\|/2$. Thus, $e_v \geq (1 - \delta)R_{\mu} \geq (1 - \delta)R_\tau \geq \frac{1 - \delta}{2} \cdot e_p$.

Final subcase 1.3.3: $f(\tau)$ is an encroached boundary $k$-simplex. We first con-
der the scenario that $f(\tau)$ is encroached by a sliver-simplex directly or indirectly.
We then know that the insertion of $p$ from $f(\tau)$ will always avoid creating
small slivers because the mesh is almost-good. Thus $R_\tau \geq bR_{f(\tau)}$. Notice that
$e_p \leq (1 + \delta)R_{f(\tau)}$. Thus, $e_v \geq (1 - \delta)bR_{f(\tau)} \geq \frac{1 - \delta}{1 + \delta}b \cdot e_p$. We then consider the
scenario that $f(\tau)$ is encroached by a $d$-simplex with a large radius-edge ratio
directly or indirectly. Here parent $p$ is selected from the picking region of $f(\tau)$.
Let $pq$ be an edge of $\tau$. Notice that $e_p \leq \|p - q\|$. We also have $R_\tau \geq \|p - q\|/2$. Thus, $e_v \geq (1 - \delta)R_{\mu} \geq (1 - \delta)R_\tau \geq \frac{1 - \delta}{2} \cdot e_p$.

Then we consider that $v$ is selected from the picking region of a boundary
$k$-simplex $\chi$, which is encroached by a $m$-simplex $\mu$.

Case 2.1: $\mu$ is encroached by a sliver-simplex directly or indirectly. Here
parent $p$ is always the circumcenter of $\mu$. Notice that $R_{\mu} \leq \sqrt{2}R_\chi$ and $e_p \leq
(1 + \delta)R_\mu$. Thus, $e_v \geq (1 - \delta)R_\chi \geq \frac{1 - \delta}{\sqrt{2}(1 + \delta)} \cdot e_p$. 
Case 2.2: \( \mu \) is encroached directly or indirectly by a \( d \)-simplex with a large radius-edge ratio. Parent \( p \) is the circumcenter of \( \mu \). We always have \( R_{\mu} \leq \sqrt{2}R_{\chi} \) and \( e_p = R_\mu \). Thus, \( e_v = R_\chi \geq R_\mu \sqrt{2} \geq \frac{\sqrt{2}}{2} \cdot e_p \).

For a vertex \( v \), as [10], we define \( D_v = \frac{lfs(v)}{e_v} \). We call \( D_v \) the density ratio of \( v \). Clearly, initially \( D_v \) is at most one for an input vertex \( v \), and after inserting new vertices, \( D_v \) tends to become larger.

**Lemma 12.** Let \( v \) be a vertex with a parent \( p = p(v) \) if there is any. Assume that \( e_v \geq C \cdot e_p \). If \( v \) is inserted due to eliminating sliver-simplex, then \( D_v \leq \frac{1+\delta}{1-\delta} + \frac{D_p}{C} \).

If \( v \) is inserted because of eliminating a \( d \)-simplex with a large radius-edge ratio, then \( D_v \leq 1 + \frac{D_p}{C \mu} \).

The proof is omitted, which is almost the same as the 3-dimensional counterpart.

**Theorem 3.** There are fixed constants \( D_k \geq 1 \), \( 1 \leq k \leq d \) such that for any vertex \( v \) inserted or rejected at the picking region of a bad \( k \)-simplex, \( D_v \leq D_k \).

Specifically, the values of \( D_i \) should satisfy the following conditions:

\[
D_d \geq \max\left\{ \frac{\varrho_0 A + \sqrt{2}^{d+1} B_{d-2}}{\varrho_0 (1-\delta) - 2d \alpha^d}, \frac{ab + \alpha^2 + \alpha B_{d-2}}{b - \sqrt{2}^{d-1} \alpha^d}, \alpha + \frac{2}{1-\delta} \right\},
\]

\[
D_{d-k} = B_{k-1} + \sqrt{2}^k \alpha^k D_d,
\]

where \( 1 \leq k \leq d-1 \), \( A = \alpha (1-\delta) + \sqrt{2}^{d+1} \), and \( B_i = \sum_{j=1}^i (\sqrt{2}^j \alpha^j) \). Hence, there is a constant \( D = \max_{k=1}^d \{D_k\} \) such that \( D_v \leq D \) for all mesh vertex \( v \).

The proof, which is based on induction, of the theorem is omitted due to space limit. The following theorem concludes that the generated mesh has good grading, i.e., for any mesh vertex \( v \), \( N(v) \) is at least some constant factor of \( lfs(v) \).

**Theorem 4.** For any mesh vertex \( v \) generated by refinement algorithm, the distance connected to its nearest neighbor vertex \( u \) is at least \( \frac{lfs(v)}{D+1} \).

The proof is omitted here, which is the same as the three-dimensional counterpart. Thus, if \( \varrho_0 > 2^d \alpha^d \) and \( b > \sqrt{2}^d \alpha^d \), our algorithm generates well-shaped Delaunay meshes with good grading. Ruppert showed that the nearest neighbor value \( N(v) \) of a mesh vertex \( v \) of any almost-good mesh is at most a constant factor of \( lfs(v) \), where the constant depends on the radius-edge ratio of the mesh. The above Lemma 12 shows that the nearest neighbor \( N(v) \) for the well-shaped Delaunay mesh is at least some constant factor of \( lfs(v) \). Notice that the number of vertices of an almost-good mesh is \( O(\int_{x \in \Omega} \frac{1}{N(x)^d} dx) \). Then we have the following theorem.

**Theorem 5.** Given a \( d \)-dimensional almost-good mesh with \( n \) vertices, the generated well-shaped mesh has \( O(n) \) vertices, where the constant depends on \( d \) and the radius-edge ratios of the meshes.
7 Discussions

In this paper, we present a refinement-based method that guarantees to remove all slivers in a $d$-dimensional almost-good simplicial mesh. Notice that the $\sigma_0$ derived from all the proofs may be too small for any practical use. We would like to conduct some experiments to see what $\sigma_0$ can guarantee that there is no small slivers created. In addition, we could have different definitions about slivers depending upon the location of the simplex: inside or near the domain boundary. This could also improve the bound on $\sigma_0$. Let $\rho_d$ be the minimum radius-edge ratio of a $d$-simplex. It is easy to show that $\rho_2 = \frac{\sqrt{3}}{3}$ and $\rho_d = \frac{1}{2\sqrt{1-\rho_{d-1}^2}}$, which is much less than the radius-edge ratio bound $\varrho_0$ achieved by our algorithm (and also the Delaunay refinement). We would like to know if we can get better radius-edge ratio bound on the $d$-simplices of the generated mesh.

References

Towards Compatible Triangulations

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Abstract. We state the following conjecture: any two planar n-point sets (that agree on the number of convex hull points) can be triangulated in a compatible manner, i.e., such that the resulting two planar graphs are isomorphic. The conjecture is proved true for point sets with at most three interior points. We further exhibit a class of point sets which can be triangulated compatibly with any other set (that satisfies the obvious size and hull restrictions). Finally, we prove that adding a small number of Steiner points (the number of interior points minus two) always allows for compatible triangulations.

1 Introduction

Can any two planar point sets (that agree on the number of points and extreme points) be triangulated in the same manner? This intuitive question is the topic of the present paper. Apart from the theoretical interest in this basic problem, questions of this kind arise in various areas of application, including image analysis, morphing, and cartography.

To define the problem more carefully, let $S_1$ and $S_2$ be two finite sets of points in the Euclidean plane. Two triangulations $T_1$ of $S_1$ and $T_2$ of $S_2$ are called compatible if the face lattices formed by their triangles, edges, and vertices (points) are isomorphic. Compatible triangulations obviously exhibit isomorphic graphs of edges. The converse is not true, in general, as an edge isomorphism between $T_1$ and $T_2$ need not preserve the external face. This led us to prefer the notion of compatible triangulations to isomorphic triangulations which sometimes has been used in the literature.

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The problem of triangulating two given point sets compatibly comes in two flavors, namely where the bijection between the points of $S_1$ and $S_2$ is either fixed or variable. The case of fixed correspondence is a known problem and has been studied in Saalfeld [7]. He pointed out that compatible triangulations do not always exist and proposed several heuristic approaches for their construction. The challenging problem left open is to determine the complexity of the related decision problem, which so far is neither known to be NP-complete nor to be polynomially solvable.

The problem becomes easier if $S_1$ and $S_2$ represent the cyclic order of the vertices of two simple $n$-gons. Given a bijection between the polygon vertices (there are only $n$ cyclic shifts to choose from) the existence can be decided in time $O(n^2)$, and compatible triangulability of the polygons can always be achieved by adding $O(n^2)$ extra points (so-called Steiner points) in either polygon; see Aronov et al. [2]. Improvements which take into account the number of reflex vertices [5] or the quality of the angles of the produced triangles [3] exist.

The present work is concerned with the problem of finding compatible triangulations without fixed point correspondence. To our knowledge this problem has not been studied before. The main question is, of course, under which propositions compatible triangulations do exist. It is clear that in both sets $S_1$ and $S_2$ the number $n$ of points must be the same, as has to be the number $h$ of points on their convex hulls $CH(S_1)$ and $CH(S_2)$, by the well-known formula $e = 3n - h - 3$ on the number of edges of a triangulation. Moreover, convex hull points necessarily map to convex hull points, and their cyclic order is preserved. This follows from the easily proven fact below: if two triangulations are compatible then any two corresponding triangles $ijk$ and $i'j'k'$ must have the same orientation, either clockwise or counter-clockwise. (We ignore the symmetric case where the orientation is reversed for each matching pair of triangles.) Intuitively speaking, this means that the underlying point sets $S_1$ and $S_2$ ‘locally’ exhibit the same order type, as defined in Goodman and Pollack [4]. It is interesting to note that the last two properties need not hold if only graph isomorphism rather than compatibility of the triangulations is required.

We will assume throughout that all point sets considered are in general position, meaning that no three points within a set lie on a common straight line. There are simple examples of point sets not in general position that do not admit compatible triangulations even if the propositions mentioned above are all met; see e.g. Krasser [6].

Throughout, let $|CH(S)|$ denote the number of extreme points in a point set $S$. Our main conjecture can be formulated in the following way.

**Conjecture 1 (Main Conjecture on Compatible Triangulations)**
Let $S_1$ and $S_2$ be two sets of points without fixed correspondence. Then compatible triangulations for $S_1$ and $S_2$ exist if these sets fulfill the following properties:

- $|S_1| = |S_2|$, 
- $|CH(S_1)| = |CH(S_2)|$, and 
- $S_1$ and $S_2$ are in general position.
Towards Compatible Triangulations

We also consider an extended version of the compatible triangulation problem, where the correspondence between the points on the convex hulls (but not between the interior points) is prescribed. The status of the extended problem is still unresolved, too.

**Conjecture 2** *(Extended Conjecture on Compatible Triangulations)*

Conjecture 1 still holds if the correspondence between the extreme points of $S_1$ and $S_2$ is given by a cyclically shifted labeling.

The motivation for stating the extended conjecture stems from the attempt to prove Conjecture 1 by induction, which naturally leads to assuming the stronger version as a hypothesis.

In this paper, first steps towards proving the conjectures above are undertaken. In fact, obtaining an affirmative answer would be a deep theoretical result, showing that all planar point sets of the same size and hull size are ‘topologically equivalent’ in this sense. The result, if true, gains in importance in view of the huge number of inequivalent order types for $n$ points; see Aichholzer et al. [1] for the exact numbers of order types for $n \leq 10$.

Section 2 exhibits a family of point sets each of whose members allows for a compatible triangulation with any other point set (within the evident size and hull restrictions). In Section 3 we prove Conjecture 2 provided the number of interior points in either set is three or less. These restricted cases already reflect part of the intrinsic complexity of the problem: the orientation of the three interior points need to be reversed to allow for a solution in some cases. Section 4 shows that compatible triangulability can always be achieved by adding at most $k - 2$ Steiner points, where $k$ denotes the number of interior points. This contrasts the afore-mentioned situation for simple polygons [2], where a quadratic number of Steiner points is necessary in the worst case, and is a result of practical relevance. Finally, Section 5 concludes the paper with some related remarks.

**2 A Family of Universal Sets**

A point set $U$ is called weak universal if for every point set $S$ with $|S| = |U|$ and $|CH(S)| = |CH(U)|$ there is a triangulation of $S$ compatible with a triangulation of $U$. The set $U$ is called universal if, in addition, the correspondence between the extreme points of $S$ and $U$ may be prescribed by a cyclic shift. Conjecture 2 if true, states that all point sets are universal. Here we prove that certain point sets are indeed universal.

Let $U$ be a point set with extreme points $u_1, \ldots, u_h$ (as they appear counterclockwise on the convex hull) and internal points $u_{h+1}, \ldots, u_n$ such that, for some $1 \leq m \leq h - 2$,

- the subset $U \setminus \{u_1, \ldots, u_m\}$ is in convex position, and
- the internal set $\{u_{h+1}, \ldots, u_n\}$ lies to the left of both $u_h u_m$ and $u_1 u_{m+1}$; see Figure 1.
We are going to prove that \( U \) is a universal set. The proof of the result requires a preliminary lemma.

**Lemma 1** Let \( P \) be a convex polygon with vertices \( v_1, \ldots, v_h \) in counterclockwise order, and let \( I \) be an arbitrary point set interior to \( P \). For any choice of indices \( 1 \leq i, j \leq h \) with \( i \leq j - 2 \) there exists a simple path \( \pi \) of edges such that

(a) path \( \pi \) connects \( v_i, I, \) and \( v_j \) (in this order) and these are the only vertices of \( \pi \);

(b) the simple polygon bounded by \( \pi \) and \( v_i v_{i+1} \ldots v_j \) can be triangulated without using an edge that connects two vertices of \( \pi \).

**Proof.** By induction on \( d = j - i \geq 2 \), the distance between \( v_i \) and \( v_j \) on \( P \)'s boundary. For \( d = 2 \) the result is trivial: sort \( I \cup \{v_i, v_j\} \) radially around \( v_{i+1} \), to produce a path star-shaped as seen from \( v_{i+1} \), and triangulate by connecting all points in \( I \) to this vertex.

Let now \( d > 2 \). If the set \( I \) is totally to the right of \( v_{i+2}v_j \) then connect \( v_i \) and \( v_{i+2} \) and slice off \( P \) the triangle \( v_i v_{i+1}v_{i+2} \) which cannot contain any point in \( I \). Apply induction to the remaining convex polygon; the boundary distance between \( v_i \) and \( v_j \) is now \( d - 1 \).

If, on the other hand, there are points of \( I \) to the left of \( v_{i+2}v_j \) then let \( p \) be the one maximizing the angle \( v_i v_{i+1}p \); see Figure 2. Connect \( p \) to \( v_{i+1} \) and \( v_{i+2} \) and observe that the resulting triangle encloses no points of \( I \). Now consider the convex polygon \( Q = pv_{i+2} \ldots v_j \) (shaded in Figure 2). Let \( I' \) and \( I'' \) be the subsets of \( I \) exterior and interior to \( Q \), respectively. By construction, no point of \( I' \) lies to the right of \( v_{i+1}p \). So, as done for the case \( d = 2 \) above, construct the path \( \pi' \) for \( I' \cup \{v_i, p\} \) that is star-shaped as seen from \( v_{i+1} \), and build the corresponding triangulation. Finally, apply induction to the polygon \( Q \) whose boundary distance between \( p \) and \( v_j \) is \( d - 1 \). This yields a simple path \( \pi'' \) for
Towards Compatible Triangulations

$I'' \cup \{p, v_j\}$ whose concatenation with \(\pi'\) obviously is still simple and satisfies the statement of the lemma. \(\square\)

**Fig. 2.** Constructing the triangulation \(T(\pi)\).

**Theorem 1** Let \(U\) be a set as described above. Then \(U\) is a universal set.

**Proof.** Let \(S\) be any point set with \(|S| = |U| = n\) and with extreme points \(v_1, \ldots, v_h\) in counter-clockwise order. We construct for \(CH(S)\) the path \(\pi\) from \(v_h\) to \(v_{m+1}\), and the associated triangulation \(T(\pi)\) on \(\pi\)'s right-hand side, whose existence is given by Lemma 1 (Note that the index \(m\) is given from \(U\) and that the counter-clockwise boundary distance between \(v_h\) and \(v_{m+1}\) is at least two.) The crucial property is that \(T(\pi)\) can be reproduced on the simple polygon \(u_h \ldots u_n u_{m+1} u_m \ldots u_1\) only edges between non-consecutive vertices of the path \(u_h \ldots u_n u_{m+1}\) lie outside this polygon, but those correspond to edges outruled by the definition of \(T(\pi)\). It remains to be observed that any triangulation of \(CH(S)\) on \(\pi\)'s left-hand side can be put on top of the convex polygon \(u_{m+1} u_{m+2} \ldots u_n\), making compatible the global triangulations. \(\square\)

**3 Few Interior Points**

Below we prove Conjecture 2 for point sets with at most three interior points. Even in these seemingly simple situations, the illusion of a quick proof is destroyed by the fact that the orientation of the three interior points might have to be reversed to achieve compatibility. We first focus on the following class of point sets.
Lemma 2. Let $S$ be a point set containing an extreme point $p$ such that $S \setminus \{p\}$ is in convex position. Then Conjecture 2 is true.

Proof. The set $S$ fulfills the requirements for the set $U$ in Section 2 for $m = 1$ and thus is a universal set by Theorem 1.

Theorem 2. Let $S_1$ and $S_2$ be two $n$-point sets each of which contains $k \leq 3$ interior points. Then Conjecture 2 is true.

Proof. The cases $k = 0$, where $S_1$ and $S_2$ are in convex position, and $k = 1$, where star-like compatible triangulations always exist, are trivial.

The case $k = 2$ is still easy. It is always possible to find two extreme points of $S_1$ labeled $a$ and $b$ such that (1) $a$ and $b$ lie on different sides of the line through the two interior points of $S_1$, and (2) the (extreme) points of $S_2$ labeled $a$ and $b$ lie on different sides of the line through the two interior points of $S_2$. (The non-existence of such a pair $a,b$ would imply that one of $S_1$ and $S_2$ totally lies on one side of the respective line.) We build two triangles for $S_1$ by joining its interior points $x$ and $y$ and connecting $a$ and $b$ to both of them. The remaining extreme points can be connected to either $x$ or $y$ accordingly. Doing the same for $S_2$ results in compatible triangulations.

The case $k = 3$ is more intriguing. Let us refer to the triangle formed by the three interior points of $S_1$ (or $S_2$) as the interior triangle of $S_1$ (or $S_2$). We distinguish between two situations, the former being similar to the case $k = 2$.

(a) $S_1$ as well as $S_2$ contains extreme points with labels $a$, $b$, $c$ each of which allows for a triangle adjacent to the respective interior triangle such that no overlap of these four triangles occurs. In this case, we integrate these triangles for both sets. For extreme points with labels intermediate to $a$, $b$, and $c$, edges can be drawn now to a unique interior point, in a compatible manner and much like the case $k = 2$.

(b) Otherwise. The solution we are going to construct will not contain the interior triangle. (In fact, no such solution exists.) Let $xyz$ be this triangle for $S_1$. Let $D_x$ be the subset of $S_1$ not on $x$’s side of the line through $y$ and $z$. Define subsets $D_y$ and $D_z$ analogously; see Figure 3(a). Since labels as above cannot be found, the subsets of $S_2$ that correspond to $D_x$, $D_y$, and $D_z$ have to look as in Figure 3(b), where two of them are totally contained in the same trilateral cell. Without loss of generality, let the third subset correspond to $D_z$. We wish to redefine $D_x$, $D_y$, and $D_z$ so as to be disjoint sets, by introducing ‘breakpoint’ labels $a$, $b$, and $c$ such that $D_x$ ranges from $a + 1$ to $b$, $D_y$ ranges from $b + 1$ to $c$, and $D_z$ ranges from $c + 1$ to $a$. This is always possible unless two of the sets are singletons consisting of the same point $p$. In this case, however, removal of $p$ leaves the set $S_1 \setminus \{p\}$ in convex position such that compatible triangulations exist by Lemma 2. Otherwise, we assign labels $x$, $y$, and $z$ to the interior triangle of $S_2$ and draw compatible edges as shown in Figure 4. Note that the orientation of $xyz$ has been reversed for $S_2$. The two shaded polygons for $S_1$ are convex and thus can be triangulated compatibly with their counterparts for $S_2$. It remains to partition the still untriangulated area for $S_1$ into convex parts. Possible reflex
Fig. 3. Two subsets move into the same cell.

Fig. 4. Compatible edge skeletons that yield convex parts.

angles may occur at $x$ or $y$. If, as in Figure 4 this happens for $x$ then there must exist a point labeled $d$ in $D_z \setminus D_y$ which makes this angle convex. (If $d$ would not exist then all extreme points of $S_2$ would lie on the same side of the line through $x$ and $z$.) The situation for $y$ is similar. This completes the construction of compatible triangulations.

4 When Steiner Points Are Allowed

Since it is still left open whether Conjectures 1 and 2 are true, the question of whether compatibility can be always forced by adding a reasonably small number of Steiner points suggests itself. We demonstrate below that a linear number suffices. An efficient algorithm for finding compatible triangulations is implicit in the proof.
Let $v$ be an extreme point of the set $S$. The ear $E(v)$ of $v$ is the triangle spanned by the two edges of $CH(S)$ incident to $v$. An ear is called empty if it does not enclose any (interior) point of $S$.

**Theorem 3** Let $S_1$ and $S_2$ be two $n$-point sets whose $n - k$ extreme points are in fixed correspondence. If $k \geq 2$ then $S_1$ and $S_2$ can be triangulated compatibly by introducing at most $k - 2$ Steiner points.

**Proof.** By induction on $k$, the number of interior points. Let $\sigma(k)$ be the number of Steiner points needed. We trivially have $\sigma(2) = \sigma(3) = 0$ by Theorem 2, so let us assume $k \geq 4$.

Let $v_1, \ldots, v_h$ be the extreme points of $S_1$, and $w_1, \ldots, w_h$ the corresponding extreme points of $S_2$, in counter-clockwise order. We may assume that for no index $i$ the ears $E(v_i)$ and $E(w_i)$ are simultaneously empty. Otherwise, compatible edges $v_{i-1}v_{i+1}$ and $w_{i-1}w_{i+1}$ are introduced, and the sets $S_1 \setminus \{v_i\}$ and $S_2 \setminus \{w_i\}$ are considered. It remains to discuss the cases below.

Case (a) There is an index $i$ such that both ears $E(v_i)$ and $E(w_i)$ are non-empty. Let $p \in S_1$ be the first point hit when rotating the edge $v_{i-1}v_i$ counter-clockwise around $v_{i-1}$. As $E(v_i)$ contains points, $p$ must be an interior point. So by introducing one Steiner point $s$ slightly to the left of $v_{i-1}p$ and close to edge $v_iw_{i+1}$, it is possible to create an empty triangle $v_ips$ and three adjacent empty triangles as shown in Figure 4(a). Obviously, the same construction can be repeated for the ear $E(w_i)$ using a suitable Steiner point $t$, giving four compatible triangles. We now apply induction to the sets $S_1 \cup \{s\} \setminus \{v_i\}$ and $S_2 \cup \{t\} \setminus \{w_i\}$ whose number of interior points is $k - 1 \geq 3$. This implies $\sigma(k) = \sigma(k-1) + 1 \leq (k - 1) - 2 + 1 = k - 2$.

Case (b) For each index $i$, exactly one of the ears $E(v_i)$ and $E(w_i)$ is empty. Clearly $S_1$ (and $S_2$) has $h \geq 4$ extreme points in this case. Let us assume $h \geq 5$ first. Then one of the sets, $S_1$ say, has at least three empty ears. Choose two of them which do not overlap, $E(v_i)$ and $E(v_j)$, for $j \geq i + 2$. Consider the convex polygon $CH(S_1) \setminus (E(v_i) \cup E(v_j))$, and some triangulation $T_1$ thereof which uses only edges which separate $E(v_i)$ and $E(v_j)$. Let $T_2$ denote the compatible triangulation of the corresponding polygon for $S_2$. Now sweep $T_1$, triangle by triangle, with a line $\ell_1$ starting at position $v_{i+1}v_{i-1}$ and stopping at position $v_{j-1}v_{j+1}$. When sweeping a triangle $v_av_bw_{b+1}$ we require $\ell_1$ to pass through point $v_a$ and to cross the edge $v_bw_{b+1}$. Sweep $T_2$ with a line $\ell_2$ in a similar manner. Let $\alpha_1$ be the current number of interior points of $S_1$ to the left of $\ell_1$ (same for $\alpha_2$). Since $E(v_i)$ and $E(v_j)$ are empty ears we have $\alpha_1 = 0$ in the beginning and $\alpha_1 = k$ at the end. On the other hand, $\alpha_2 \geq 1$ in the beginning and $\alpha_2 \leq k - 1$ at the end, because $E(w_i)$ and $E(w_j)$ are both non-empty. As a consequence, there exist indices $a$ and $b$ that yield $\alpha_1 = \alpha_2$ while $\ell_1$ sweeps the triangle $v_av_bw_{b+1}$ and $\ell_2$ sweeps the corresponding triangle $w_aw_bw_{b+1}$. Now pick a Steiner point $s$ on (this very position of) $\ell_1$ and close to $v_bw_{b+1}$, and put in edges from $s$ to $v_b$, $v_{b+1}$, and $v_a$; see Figure 5(b). Repeat the construction for $S_2$ which gives a compatible triangle. Finally, apply induction to the respective subsets of $S_1$ and $S_2$ on either side of the lines. They have $k'$ and $k''$ interior
points, respectively, for $1 \leq k', k'' \leq k - 1$. We may assume $k' \geq 2$ because $k \geq 4$. We get $\sigma(k) \leq \sigma(k') + \sigma(k'') + 1$, with $\sigma(k') \leq k' - 2$ and $\sigma(k'') \leq k'' - 1$ and equality in the last term if $k'' = 1$. In summary, $\sigma(k) \leq k - 2$ again.

The case $h = 4$ which has not been covered yet is easier. There is a line $\ell_1$ crossing the edges $v_1v_2$ and $v_3v_4$ which has exactly 2 interior points of $S_1$ to its left and $k - 2 \geq 2$ such points to its right. There is also a similar line $\ell_2$ that crosses the corresponding edges for $S_2$. Pick two Steiner points on $\ell_1$, one close to $v_1v_2$ and the other close to $v_3v_4$, and create the two respective triangles, as shown in Figure 5(c). Repeat the construction for $\ell_2$ and $S_2$ which results in two compatible triangles. Now apply induction to the respective subsets of $S_1$ and $S_2$ on either side of the lines. This gives $\sigma(k) \leq \sigma(2) + \sigma(k - 2) + 2 \leq (k - 4) + 2 = k - 2$ which completes the argumentation.

5 Remarks

It is a challenging open problem to prove our conjectures for general point sets. Unfortunately, the constructions in the proof of Theorem 2 do not seem to generalize easily to the case of more interior points. Approaches to prove Conjecture 2 by induction suffer from the lack of appropriate methods for splitting the problem. Here is a discouraging related fact: given two quadrilaterals, one convex and the other star-shaped, plus one interior point for each, it may not be possible to triangulate them compatibly if the correspondence between their boundary vertices is fixed by a cyclic shift.

We tested Conjecture 2 for small point set sizes, utilizing a database of combinatorially inequivalent point sets (i.e., order types); see [1]. If two point sets exhibit the same order type then every triangulation of one set also leads to a compatible triangulation for the other. By checking all different pairs of order types we could verify the conjecture for up to 8 points. Although our database provides all order types of size up to 10, exhaustive tests turned out to be time
consuming and are still incomplete for size 9. In fact, determining the time complexity of computing compatible triangulations, if they exist, for two given $n$-point sets is an open problem.

Finally, we pose the problem of triangulating compatibly two point sets in fixed correspondence when Steiner points are allowed. In view of the applications mentioned in [7], a fast algorithm using a small number of Steiner points would be of practical relevance.

References

An Improved Upper Bound on the Size of Planar Convex-Hulls

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Abstract. Let $C$ be the convex-hull of a set of points $S$ with integral coordinates in the plane. It is well–known that $|C| \leq cD^{2/3}$ for some constant $c$ where $D$ is the diameter of $S$: i.e. the maximum distance between any pair of points in $S$. It has been shown that $c = 7.559..$ for an arbitrary $S$, and $c = 3.496..$ in the special case when $S$ is a ball centered at the origin in the plane. In this paper we show that $c = 12/\sqrt[3]{4\pi^2} = 3.524..$ is sufficient for an arbitrary set of lattice points $S$ of diameter $D$ in the plane, and $|C| \sim 12\sqrt[3]{2/(9\pi^2)}D^{2/3} = 3.388..D^{2/3}$ is achieved asymptotically. Our proof is based on the construction of a special set in first quadrant, and the analysis of the result involves the calculation of the average order of certain number-theoretical functions associated with the Euler totient function $\phi(n)$.

1 Introduction

A lattice point is a point with integral coordinates. Given a set $S$ of lattice points in the plane, let $C$ be the convex-hull of $S$, and denote the number of extreme points in $C$ by $|C|$. The behavior of $|C|$ as a function of parameters associated with $S$ has been studied in various contexts in computational geometry, computer graphics [10,6,9], and integer programming [1,3,5].

Andrews’ general theorem [1] on convex bodies in $d$-dimensional space implies that in the plane $|C|$ is bounded by

$$|C| \leq cD^{2/3}$$

for some constant $c$, where $D$ is the diameter of $S$: i.e. the maximum distance between any two points in $S$. Another proof of this bound for the plane was given by Katz and Volper [9]. The constant $c$ has also been studied. Balog and Bárány [2] showed that when $S$ is the ball of radius $r = D/2$ centered at the origin, i.e. $S = rB^2$, then

$$0.207..D^{2/3} \leq |C| \leq 3.496..D^{2/3}.$$
Therefore one can take $c = 3.496..$ in the special case when $S$ is a ball of diameter $D$. A $d$-dimensional analysis for the ball appears in [4]. Har-Peled [7] showed that the value $c = 6\sqrt{2}/7.559..$ is sufficient for the bound (1) in the plane for arbitrary $S$.

In this paper we investigate further the nature of the constant $c$ in (1) and show that

$$c = 12 / 3\sqrt{2/\pi^2} = 3.524..$$

suffices for an arbitrary $S$ of diameter $D$ in the plane. Our proof is based on the construction of a special set in first quadrant of the plane which satisfies certain constraints. The construction involves selecting a set of fractions in a particular order as slopes of the line segments of the convex-hull. Based on the properties of the average order of certain number-theoretical functions associated with the Euler totient function $\phi(n)$, we derive an upper bound on the size of the set constructed. This leads to an improved value for $c$. We also show that using the construction idea of the proof, we can always create a convex-hull $C$ with a given diameter $D$ such that

$$|C| \sim 12 \sqrt{2/(9\pi^2)} D^{2/3} = 3.388..D^{2/3}$$

is achieved asymptotically.

The organization of this paper is as follows. Section 2 gives the number-theoretical background we require, and includes the proof of the main theorem, which we then use in Section 3 for proving our result on the improved value of $c$. In Section 4 we construct large convex-hulls with a given diameter proving (2).

2 Number-Theoretical Definitions and Results

We use the classical book by Hardy and Wright [8] as our main reference for the definitions and basic results used in this section. We denote by $(r, s)$ the greatest common divisor of $r$ and $s$.

- The Euler totient function $\phi(n)$ ([8], p. 52) is defined as follows:
  1. $\phi(1) = 1$
  2. for $n > 1$, $\phi(n)$ is the number of positive integers less than $n$ and relatively prime to $n$.

- The Möbius function $\mu(n)$ ([8], p. 234) is defined by
  1. $\mu(1) = 1$,
  2. $\mu(n) = 0$ if $n$ has a square factor,
  3. $\mu(p_1 p_2 \cdots p_k) = (-1)^k$ if all the primes $p_1, p_2, \ldots, p_k$ are different.

- The Riemann zeta function $\zeta(s)$ ([8], p. 245) is defined for $s > 1$ by

$$\zeta(s) = \sum_{n=1}^{\infty} \frac{1}{n^s}.$$ 

We use the following well-known results:
Lemma 1. 

\[ \phi(n) = \sum_{d|n} \frac{n}{d} \mu(d), \quad (3) \]

\[ \zeta(2)^{-1} = \sum_{d=1}^{\infty} \frac{\mu(d)}{d^2} = \frac{6}{\pi^2}. \quad (4) \]

Proof. The proof of (3) is in (8), p. 235. The proof of (4) can be found in (8), Thm. 287, p. 250; Thm. 293, p. 251.

Let \( \Phi_0(n) = \phi(1) + \phi(2) + \cdots + \phi(n) \). It is known in relation to Farey fractions (8, Thm. 330, p. 268) that

\[ \Phi_0(n) = \frac{3n^2}{\pi^2} + O(n \log n), \quad (5) \]

so that the average order ((8, p. 263) of the function \( \phi(n) \) is given by (5).

Theorem 1 below gives an expression for the average order of the function \( n^r \phi(n) \) for \( r \geq 0 \), generalizing (5).

**Theorem 1.** For any integer \( r \geq 0 \)

\[ \Phi_r(n) = 1^r \cdot \phi(1) + 2^r \cdot \phi(2) + \cdots + n^r \cdot \phi(n) = \frac{6}{(r + 2)\pi^2} n^{r+2} + O(n^{r+1} \log n). \quad (6) \]

Proof. By (3) of lemma 11 we have

\[ \Phi_r(n) = \sum_{m=1}^{n} m^r \phi(m) = \sum_{m=1}^{n} m^r \sum_{d|m} \frac{m}{d} \mu(d) = \sum_{m=1}^{n} \sum_{dd'=m} m^r d' \mu(d) \]

\[ = \sum_{m=1}^{n} \sum_{dd'=m} d'^r (d')^{r+1} \mu(d) = \sum_{dd' \leq n} d'^r (d')^{r+1} \mu(d) = \sum_{d=1}^{n} d^r \mu(d) \sum_{d'=1}^{\lfloor n/d \rfloor} (d')^{r+1} \]

Using the fact that

\[ \sum_{k=1}^{n} k^{r+1} = \frac{n^{r+2}}{r + 2} + O(n^{r+1}), \]

we get

\[ \sum_{d'=1}^{\lfloor n/d \rfloor} (d')^{r+1} = \frac{1}{r + 2} \left( \frac{n}{d} \right)^{r+2} + O \left( \frac{n^{r+1}}{d^{r+1}} \right). \]

Therefore

\[ \Phi_r(n) = \frac{n^{r+2}}{r + 2} \sum_{d=1}^{n} \frac{\mu(d)}{d^2} + O \left( n^{r+1} \sum_{d=1}^{n} \frac{1}{d} \right) \]

\[ = \frac{n^{r+2}}{r + 2} \sum_{d=1}^{\infty} \frac{\mu(d)}{d^2} + O \left( n^{r+2} \sum_{d=n+1}^{\infty} \frac{1}{d^2} \right) + O \left( n^{r+1} \log n \right) \]

\[ = \frac{n^{r+2}}{r + 2} \zeta(2)^{-1} + O(n^{r+1}) + O(n^{r+1} \log n), \]
and the theorem follows from the last equality and (4) of lemma 1.

We need the following result relating $\Phi_0(n)$ and $\Phi_1(n)$ as a step in our study of $|\mathcal{C}|$.

**Theorem 2.** Let $\Phi_0(n) = \phi(1) + \phi(2) + \cdots + \phi(n)$,
$$\Phi_1(n) = 1 \cdot \phi(1) + 2 \cdot \phi(2) + \cdots + n \cdot \phi(n).$$

Then
$$\Phi_0(n) \sim \frac{3}{\sqrt{4\pi^2}} \Phi_1(n)^{2/3} = 0.8810516 \cdot \Phi_1(n)^{2/3}.$$ (7)

**Proof.** The theorem follows by combining the expressions for $\Phi_0(n)$ and $\Phi_1(n)$ obtained as the cases $r = 0$, and $r = 1$ of theorem 1.

We note that also the magnitude of the error term in (7) can be calculated by using the full expressions for $\Phi_0(n)$ and $\Phi_1(n)$ obtained. This gives
$$\Phi_0(n) = \frac{3}{\sqrt{4\pi^2}} \Phi_1(n)^{2/3} + O \left( \Phi_1(n)^{1/3} \log \Phi_1(n) \right).$$

We omit the details of this calculation.

### 3 An Improved Upper Bound

We first establish an upper bound on the size of the convex-hull of lattice points in first quadrant of the plane in the following lemma.

**Lemma 2.** Let $S_{a,b}$ be a set of lattice points in first quadrant of the $xy$-plane enclosed by $y = a$ and $x = b$. For the convex-hull $\mathcal{C}_{a,b}$ of $S_{a,b}$, there holds
$$|\mathcal{C}_{a,b}| \leq \frac{3}{\sqrt{4\pi^2}} (a + b)^{2/3} + O \left( (a + b)^{1/3} \log (a + b) \right).$$

**Proof.** Without loss of generality, $S_{a,b}$ includes the points $(0, a)$ and $(b, 0)$ and both $a$ and $b$ are positive integers (Figure 1). Instead of the point-set $\mathcal{C}_{a,b}$ we consider the corresponding set of slopes $R_{a,b}$ of the line segments (edges) connecting the consecutive extreme points of the convex-hull. Clearly, $|\mathcal{C}_{a,b}| = |R_{a,b}| + 1$, so we can alternately study the properties of the set $R_{a,b}$. By convexity, the slopes of the edges of the convex-hull are all different. Furthermore $S_{a,b}$ is a set of lattice points, and therefore the slopes of the non-vertical edges are all rational numbers.

Let $R_{a+b}^*$ be an optimal set for the following problem: maximize $|R|$ subject to
$$\sum_{\frac{y}{b} \in R} y + \sum_{\frac{x}{a} \in R} x = \sum_{\frac{y}{b} \in R} y + x \leq a + b$$ (8)

Clearly, for any $a$ and $b$, $|R_{a,b}| \leq |R_{a+b}^*|$. We will find a bound for $|R_{a+b}^*|$. Let $Q_i$ for $i \geq 0$ be the set of slopes defined as follows:
1. \( Q_0 = \emptyset \),
2. \( Q_1 = \{ \left( \frac{0}{1}, \frac{1}{0} \right) \} \). We assume that the fraction \( \frac{1}{0} \) is defined as a slope and it represents the vertical edge whose length is one unit.
3. For \( i > 1 \), \( Q_i = \{ \left( \frac{y}{x} \right) \mid y + x = i \text{ and } (y, x) = 1 \} \).

Table I illustrates the first few values \( Q_i, \phi(i), \Phi_0(i), \text{ and } \Phi_1(i) \). The following properties can easily be seen for \( i > 1 \):

\[
\begin{align*}
|Q_i| &= \phi(i) \quad (9) \\
\sum_{\frac{y}{x} \in Q_i} y + x &= i |Q_i| \quad (10)
\end{align*}
\]

Let \( n + 1 \) be the smallest number such that \( \sum_{i=1}^{n+1} i |Q_i| > a + b \), and \( Q'_{n+1} \) be an arbitrary subset of \( Q_{n+1} \) such that

\[
|Q'_{n+1}| = \left\lfloor \frac{(a + b) - \sum_{i=1}^{n} i |Q_i|}{n + 1} \right\rfloor
\]

Then consider the following set

\[
R_{a+b} = \left( \bigcup_{i=1}^{n} Q_i \right) \cup Q'_{n+1}.
\]

We claim that \( R_{a+b} \) is a maximal set which satisfies the constraint \([\text{S}]\). The expression \([\text{II}]\) describes a greedy construction: To include in set \( R_{a+b} \), select a fraction whose numerator-denominator sum is the smallest. Continue including fractions until the sum of all the numerators and the denominators of the fractions currently in the set exceeds \( a + b \). Since \( R_{a+b} \) is a set, no fraction can be included in the set more than once, and since the slopes are all different, among the equivalent fractions, the irreducible one has the smallest numerator-denominator sum. This explains why only the relatively prime numbers are to be considered in the construction.
Table 1. Table of $Q_i$, $\phi(i)$, $\Phi_0(i)$, and $\Phi_1(i)$ for $i = 1, 2, \ldots, 7.$

<table>
<thead>
<tr>
<th>$i$</th>
<th>$Q_i$</th>
<th>$\phi(i)$</th>
<th>$\Phi_0(i)$</th>
<th>$\Phi_1(i)$</th>
</tr>
</thead>
<tbody>
<tr>
<td>1</td>
<td>${0, 1/6}$</td>
<td>1</td>
<td>1</td>
<td>1</td>
</tr>
<tr>
<td>2</td>
<td>${1}$</td>
<td>1</td>
<td>2</td>
<td>3</td>
</tr>
<tr>
<td>3</td>
<td>${1/2, 2/3}$</td>
<td>2</td>
<td>4</td>
<td>9</td>
</tr>
<tr>
<td>4</td>
<td>${1/3, 3/4}$</td>
<td>2</td>
<td>6</td>
<td>17</td>
</tr>
<tr>
<td>5</td>
<td>${1/4, 2/3, 3/4, 4/5}$</td>
<td>4</td>
<td>10</td>
<td>37</td>
</tr>
<tr>
<td>6</td>
<td>${1/5, 5/6}$</td>
<td>2</td>
<td>12</td>
<td>49</td>
</tr>
<tr>
<td>7</td>
<td>${1/6, 2/5, 3/4, 4/5, 5/6}$</td>
<td>6</td>
<td>18</td>
<td>91</td>
</tr>
</tbody>
</table>

From (11) we have

$$\bigcup_{i=1}^{n} Q_i \subseteq R_{a+b} \subseteq \bigcup_{i=1}^{n+1} Q_i$$  \hspace{1cm} (12)

Using (12) and (9) we find that

$$\sum_{i=1}^{n} \phi(i) + 1 \leq |R_{a+b}| \leq \sum_{i=1}^{n+1} \phi(i) + 1$$  \hspace{1cm} (13)

Again using (12) and (9) together with the expression (10), and using the fact that by construction $n+1$ is the smallest number such that $\sum_{i=1}^{n+1} i|Q_i| > a + b$ we get

$$\sum_{i=1}^{n} i\phi(i) + 1 \leq a + b \leq \sum_{i=1}^{n+1} i\phi(i) + 1$$  \hspace{1cm} (14)

We note that $\phi(n+1) = O(\Phi_1(n)^{1/3})$ using the expression $\Phi_1(n)$ obtained from (6) with $r = 1$. Therefore we can write the following upper bound for $|R_{a+b}|$ using (13):

$$|R_{a+b}| = \sum_{i=1}^{n} \phi(i) + O \left( \Phi_1(n)^{1/3} \right) = \Phi_0(n) + O \left( \Phi_1(n)^{1/3} \right)$$

Furthermore from Theorem 2

$$|R_{a+b}| = \frac{3}{\sqrt[3]{4\pi^2}} \Phi_1(n)^{2/3} + O \left( \Phi_1(n)^{1/3} \log \Phi_1(n) \right) + O \left( \Phi_1(n)^{1/3} \right),$$  \hspace{1cm} (15)

and from (14) we obtain that $\Phi_1(n) \sim a + b$. Therefore

$$|R_{a+b}| = \frac{3}{\sqrt[3]{4\pi^2}} (a + b)^{2/3} + O \left( (a + b)^{1/3} \log (a + b) \right)$$

which proves the bound for $|C_{a,b}|$ of the lemma.
Theorem 3. For the convex-hull $C$ of a set $S$ of lattice points of diameter $D$ in the plane, there holds

$$|C| \leq \frac{12}{\sqrt{4\pi^2}} \, D^{2/3} + O\left(D^{1/3} \log D\right)$$

Proof. $S$ can be enclosed by an axis-parallel rectangle whose sides are at most $D$ in length. Using this rectangle, we can partition $S$ into four parts such that each part is confined to a single quadrant as shown in Figure 2. For the size of $C$ we have

$$|C| \leq |C_{a_1,b_1}| + |C_{a_2,b_2}| + |C_{a_3,b_3}| + |C_{a_4,b_4}|$$

By Lemma 2 we can rewrite the inequality as

$$|C| \leq \frac{3}{\sqrt{4\pi^2}} \left((a_1+b_1)^{2/3} + (a_2+b_2)^{2/3} + (a_3+b_3)^{2/3} + (a_4+b_4)^{2/3}\right) + O(D^{1/3} \log D)$$

(16)

![Fig. 2. Partitions of $S$ in quadrants.](image)

The function $f(x) = x^{2/3}$ is concave on $[0, \infty)$. In particular for nonnegative $x_1$ and $x_2$,

$$x_1^{2/3} + x_2^{2/3} \leq 2 \left(\frac{x_1 + x_2}{2}\right)^{2/3}.$$

Therefore, the equality in (16) becomes

$$|C| \leq \frac{3}{\sqrt{4\pi^2}} \left(\frac{4}{4} \left(a_1 + b_1 + a_2 + b_2 + a_3 + b_3 + a_4 + b_4\right)^{2/3}\right) + O\left(D^{1/3} \log D\right)$$

Since $a_1 + b_2$, $a_3 + b_4$, $a_2 + b_3$, and $a_4 + b_1$ are all smaller than or equal to $D$, the inequality for $|C|$ of the theorem holds.
Constructing a Large Convex-Hull with a Given Diameter

To investigate how tight the new value of \( c \) in the upper bound for \( |C| \) is, we revisit the proof of Lemma 2. We first describe how we can actually construct a convex-hull \( C_{a+b} \) from the set of slopes \( R_{a+b} \) created in the proof for particular \( a \) and \( b \). More precisely, for every \( n \), we construct a convex-hull \( C(n) \) in first quadrant with \( a = b = \frac{1}{2} \Phi_1(n) + \frac{1}{2} \), symmetric about the line \( y = x \). As its set of slopes, we take \( R(n) = \bigcup_{i=1}^{n} Q_i \).

\( C(n) \) connects the point \( (0, \frac{1}{2} \Phi_1(n) + \frac{1}{2}) \) on the \( y \)-axis to \( (\frac{1}{2} \Phi_1(n) + \frac{1}{2}, 0) \) on the \( x \)-axis and consists of \( \Phi_0(n) \) edges. We start with the point \( (0, \frac{1}{2} \Phi_1(n) + \frac{1}{2}) \) as the first vertex, and if the \( i \)-th smallest fraction in \( R(n) \) is \( y_i/x_i \), then we place the next vertex on the hull \( x_i \) units to the right and \( y_i \) units down from the current vertex. See Figure 4 for the \( n = 5 \) case.

Now consider the \( C(n) \) constructed as we described. If it includes a vertex whose distance from the origin is larger than \( \frac{1}{2} \Phi_1(n) + \frac{1}{2} \), then the diameter of the set constructed by taking four copies of \( C(n) \) around the origin will have diameter larger than \( D \). That this is indeed the case can be seen from Figure 3 where the point farthest from the origin in our construction is not inside the ball with radius 46. Thus the constant \( c = \frac{12}{\sqrt{4\pi^2}} = 3.352.. \) in the upper bound is not tight.

We next investigate how close we can come to this value. It follows by symmetry and convexity that the farthest vertex from the origin on \( C(n) \) constructed is one of the endpoints of the edge with slope 1 on the hull as shown in Figure 3.
Denote by $h(n)$ the $x$-coordinate of this point, and by $r(n)$ its distance from the origin. By construction

$$ h(n) = \sum_{i+j \leq n} j \chi[(i, j) = 1] \chi[i \leq j] = \sum_{j=1}^{n} \sum_{i=1}^{j} \chi[(i, j) = 1] \chi[i + j \leq n] \quad (17) $$

where $\chi$ is the indicator function of its argument, i.e., it is 1 if its argument evaluates to true, and 0 otherwise.

We next show that $h(n) \sim \frac{3}{8} \Phi_1(n)$. Since intercepts of the hull are $a(n) = b(n) = \frac{1}{2} \Phi_1(n) + \frac{1}{2}$ this implies $h(n) \sim \frac{3}{4} a(n)$, and we can calculate $r(n)$ by the Pythagorean formula as

$$ r(n) \sim \frac{3\sqrt{2}}{4} a(n). \quad (18) $$

**Theorem 4.** Suppose $h(n)$ is as defined in [17], and $\Phi_1(n)$ as defined in Theorem 2. Then $h(n) = \frac{3}{8} \Phi_1(n) + O(n^2 \log n)$.

**Proof.** We will indicate the derivation of the asymptotic part and ignore the calculation of the error terms. First we write $h(n) = h_1(n) + h_2(n)$ where in $h_1(n)$, the index $j$ runs from 1 to $n/2$, and in $h_2$, $j$ runs from $n/2 + 1$ to $n$. Then

$$ h_1(n) = \sum_{j=1}^{n/2} j \sum_{i=1}^{j} \chi[(i, j) = 1], $$

since for the indices in question $i \leq j$ and we automatically have $i + j \leq n$. Therefore $h_1(n) = \Phi_1(\frac{n}{2})$. On the other hand, if $j \geq n/2$, then $n - j \leq j$ and

$$ h_2(n) = \sum_{j=n/2}^{n} j \sum_{i=1}^{j} \chi[(i, j) = 1] \chi[i \leq n - j] $$
Using the expressions $\Phi_0(n) \sim 3n^2 / \pi^2$, and $\Phi_1(n) \sim 2n^3 / \pi^2$ from Theorem 2, and adding the terms for $h_1(n)$ and $h_2(n)$, we obtain
\[ h(n) \sim \left( \frac{1}{4} + 3 - \frac{3}{4} - 2 + \frac{1}{4} \right) \frac{n^3}{\pi^2} = \frac{3n^3}{4\pi^2} \sim \frac{3}{8} \Phi_1(n) \]
as claimed.

**Theorem 5.** For $D$ large, there exists a set of lattice points in the plane with diameter $D$ and convex-hull $C$ such that
\[ |C| \sim 12 \sqrt[3]{\frac{2}{9\pi^2}} D^{2/3} = 3.388..D^{2/3}. \]

**Proof.** If we set $a = b = \frac{2}{3\sqrt{2}} D$ where $D$ is of the form $\Phi_1(n)$, and use the construction for $C(n)$, then $r \sim D/2$ by (18). By Lemma 2
\[ |C(n)| \sim \frac{3}{\sqrt[3]{4\pi^2}} \left( \frac{4}{3\sqrt{2}} \right)^{2/3} D^{2/3} = 3 \sqrt[3]{\frac{2}{9\pi^2}} D^{2/3}. \tag{19} \]
Since identical convex-hulls are constructed in all four quadrants to obtain $C$ from $C(n)$, the diameter of $C$ is $D$ and $|C| \sim 4|C(n)|$. This proves the claim of the theorem when combined with (19).

**References**

On the Planar Two-Watchtower Problem*

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Abstract. In this paper we study the following planar two-watchtower problem: given a terrain (X-monotone chain) $T$ with $n$ vertices, locate two vertical segments (watchtowers) $uv$ and $u'v'$ which have to be erected on $T$ such that every point on $T$ is visible to the top of either watchtowers ($u$ or $u'$) and the maximum height of $uv, u'v'$ is minimized. We present an $O(n^4)$ time algorithm to solve the discrete version of this problem when $v, v'$ must be on some vertices of $T$. Under a mild condition on solving a special cubic equation with three bounded variables in $O(f_3)$ time we can also generalize the algorithm to solve the general problem in $O(n^4 + n^3 f_3)$ time. Using parametric search, the discrete problem can be solved in $O(n^3 \log^2 n)$ time and the general problem can be solved in $O(n^4 \log^2 n)$ time.

1 Introduction

Given a polyhedral terrain (mountain landscape), the shortest watchtower problem is to compute the shortest vertical segment $uv$ erected on the terrain such that every point on the terrain can be seen from the top of $uv$, say $u$. The problem arises in geographical information systems, military surveillance, etc. The problem was first studied by Sharir [Sh88], who presented an $O(n \log^2 n)$ time algorithm (where $n$ is the size of the terrain). The algorithm was later improved by Zhu to $O(n \log n)$ time [Zh97]. On the related problems of guarding polyhedral terrains, readers are referred to [BDEG94,BSTZ97,O’R97,WZ00,Zh94].

A very natural generalization of the problem is to build a set of watchtowers which can collectively guard the terrain. The objective is certainly to minimize

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the maximum (or sum) of the heights of the watchtowers. This general problem (in 3D) is NP-hard as the problem of deciding whether a set of points (which can be thought of as watchtowers with zero height) can collectively guard a given terrain is NP-hard [CS89]. Eidenbenz et al. proved that positioning the minimum number of watchtowers with fixed height (to cover a 3D terrain) is not only NP-hard, but also NP-hard to approximate within a ratio of $C \log n$, where $C$ is a constant [ESW98]. However, if the number of towers $k (k > 1)$ is fixed then almost nothing is known, yet the problem seems to be difficult even for $k = 2$ in 3D. In this paper we first study the corresponding problem in 2D, i.e., given a terrain in 2D, compute two watchtowers $u_1v_1$ and $u_2v_2$ (with heights $|u_1v_1|$ and $|u_2v_2|$ respectively) such that the terrain can be collectively guarded by the two watchtowers and the maximum of $\{|u_1v_1|, |u_2v_2|\}$ is minimized.

Regarding guarding a 2D terrain (monotone chain), Chen et al. proved that it is NP-hard to compute the minimum number of guards (which can only lie on the 2D terrain) to cover the whole terrain [CEU95]. If the guards are all at fixed height above the 2D terrain, then Nilsson’s result implies a polynomial time solution to find the minimum number of guards [Ni94]. Our problem (even in 2D) is new and never been studied before, to the best of our knowledge. We first present some basic definitions related to this paper. Given a point $p$ in the plane, let $x(p), y(p)$ be the $x, y$-coordinate of $p$ respectively. A polygonal chain $< v_1, ..., v_n >$ is said to be $X$-monotone if $x(v_i) > x(v_j)$ implies $i > j$, for all $i$ and $j$ between 1 and $n$. A terrain $T$ in 2D is a $X$-monotone chain with vertices $v_1, ..., v_n$. Given two points $a, b$ on or above $T$, $a, b$ are visible to each other (or $a, b$ covers each other) if the segment $ab$ does not intersect any edge of $T$. The visibility region of $e \in T$ from $a$, is the set of all points of $e \in T$ which are visible to $a$. These definitions can all be found in [PS85, O’R87, O’R94].

2 Algorithm

In this section, we present implementable polynomial time algorithms for the planar 2-watchtower problem. We also consider the more theoretical solutions using parametric search. We first consider a discrete version of the problem, namely, when the base (bottom vertex) of each of the two watchtowers must be at a vertex of $T$. The following property regarding 2D visibility is easy to prove.

Lemma 1. Given any point $u$ above $T$, the non-empty visibility region of any edge $v_iv_{i+1}$ from $u$ is connected; moreover, if $x(u) < x(v_i)$ then the non-empty visibility region contains $v_{i+1}$ and if $x(u) > x(v_{i+1})$ then the non-empty visibility region contains $v_i$.

2.1 Algorithm for the Discrete Planar 2-Watchtower Problem

We begin with some definitions. Given a point $p$ on $T$, we call the minimum height watchtower located at $v_i$ whose top can see $p$ the critical watchtower of $p$ and $v_i$. We say that the height of this critical watchtower is the critical height.
of \( p \) and \( v_i \) and denote it by \( h(p,v_i) \). Clearly \( h(p,v_i) \) is the minimum height of a watchtower one can build which is located at \( v_i \) and its top can see \( p \). In Figure 1, \( \overline{u_1v_i} \) is the critical watchtower of \( p \) and \( v_i \) and its height \(|u_1v_i|\) is exactly \( h(p,v_i) \). Similarly we define the critical watchtower of an edge \( e \) of \( T \) and \( v_i \) as the minimum height watchtower located at \( v_i \) whose top can see every point of \( e \) and we denote the critical height of this tower by \( h(e,v_i) \). In Figure 1, \( \overline{u_2v_i} \) is the critical watchtower of \( e_4=v_4v_5 \) and \( v_i \), and \( h(e_4,v_i) = |u_2v_i| \). (Notice that the concept of critical watchtower can be easily generalized to arbitrary point \( v \) on \( T \).) In the following we first prove some fundamental properties (a necessary condition) of the discrete planar 2-watchtower problem.

**Lemma 2.** When \( i,j \) are fixed, if \( u_i \overline{v_i} \) (\( \overline{u_jv_j} \)) achieves the optimal solution for the discrete 2-watchtower problem then \( u_i,u_j \) collectively see \( T \) and there exists a point \( q \) on \( T \) such that \( h(q,v_j) \geq h(q,v_i) \) (\( h(q,v_i) \geq h(q,v_j) \)).

Based on this lemma, we have the following algorithm to solve the discrete 2-watchtower problem. Without loss of generality, we assume that \( i,j \) are fixed for now. We have to consider two cases: (1) \( q \) is a vertex of \( T \), \( q \neq v_1,q \neq v_n \), and (2) \( q \) is a point on some edge of \( S \). Notice that in case 1, following Lemma 1, one of the edges incident to \( q \) must be completely seen by the optimal watchtower \( \overline{u_iu_i} \) (\( \overline{u_jv_j} \)). Moreover, if \( q = v_d \), \( d < i \), then the edge \( \overline{v_{d-1}v_d} \) must be completely seen by \( u_i \). Symmetrically, if \( q = v_d \), \( d < i \), then the edge \( \overline{v_{d-1}v_d} \) must be completely seen by \( u_i \). Therefore, to handle case 1 we do the following: for each edge \( e \) we compute \( h(e,v_i) \) and \( h(e,v_j) \), if \( h(e,v_j) \geq h(e,v_i) \) we build two towers at \( v_i,v_j \) with height \( h(e,v_i) \), otherwise we build two towers at \( v_i,v_j \) with height \( h(e,v_j) \). Then we check whether the tops of the two towers, \( u_i \) and \( u_j \), can collectively see \( T \). For a fixed \( e \), this can be done in linear time: just compute the visibility region of \( T \) from \( u_i,u_j \) and then check whether every edge can be collectively covered by \( u_i \) and \( u_j \). This can be done in linear time using the standard visibility
computation subroutines [EA81,Le83,Jo90,JS87,O’R87]. We record all pairs of the towers when their tops can collectively cover $T$. Among them we pick up the pair with the minimum height as $H_1(T, i, j)$. Clearly $H_1(T, i, j)$ can be computed in $O(n^2)$ time and $O(n)$ space as there are $O(n)$ edges in $T$. We remark that $h(e, v_i)$ over all $e$ can be pre-computed in linear time with the shortest path map rooted at $v_i$ [GHLST87], but that will not affect the overall complexity for computing $H_1(T, i, j)$.

For case 2, where a point in the middle of some edge determines the optimal solution for the discrete two-watchtower problem, we do the following. We fix an edge $e$ and assume that $q$ must be on $e$ (eventually we need to try on all such $e$). We first prove the following lemma.

**Lemma 3.** When $i, j$ are fixed, in case 2, $q$ cannot be to the left (right) of both $v_i$ and $v_j$. In other words, $e$ must be between $v_i$ and $v_j$.

Now we show how to handle case 2. Without loss of generality, we assume $x(v_i) < x(v_j)$ and $e = v_kv_{k+1}$ is between $v_i$ and $v_j$. We first compute the shortest paths from $v_k$ to $v_i$ and from $v_{k+1}$ to $v_j$. Let the two paths be $<p_1 = v_k, p_2, \ldots, p_h = v_i>$ and $<q_1 = v_{k+1}, q_2, \ldots, q_m = v_j>$ respectively (Figure 2).

Notice that if we extend the edges $p_sp_{s+1}(s = 1, \ldots, h - 2)$, they naturally divide $e$ into $O(h)$ red intervals and at the same time the intersections of these extensions with the vertical line through $v_i$ determines the heights of the towers located at $v_i$. We can define a function $F_r(q)$ for each point $q$ on $e$ such that it is the height of the tower located at $v_i$ and the top of the tower can see $q$, but nothing to the left of it. Clearly $F_r(q)$ is piecewise decreasing function (containing $O(h)$ pieces). Symmetrically, we can extend the edges $q_tq_{t+1}(t = 1, \ldots, m - 2)$, they also naturally divide $e$ into $O(m)$ blue intervals. We can also define a function $F_b(q)$ for each point $q$ on $e$ such that it is the height of the tower located

![Fig. 2. Handling case 2 for the discrete two-watchtower problem.](image)
at \( v_j \) and the top of the tower located at \( v_j \) can see \( q \), but nothing to the right of it. Again \( F_b(q) \) is a piecewise increasing function (containing \( O(m) \) pieces). (In other words, \( F_r(q), F_b(q) \) are exactly \( h(q, v_i), h(q, v_j) \) respectively.) Let \((X, Y)\) be the coordinate of \( q \), \((x, y)\) be the coordinate of \( u_i \) and let \( p_s = (x(p_s), y(p_s)) \) be the vertex of \( T \) tangent to \( \overline{u_iq} \). The height of \( \overline{u_iv} \) is the difference between \( y \) and \( y(v_i) \), which is

\[
(x(v_i) - x(p_s)) \frac{Y - y(p_s)}{X - x(p_s)} + y(p_s) - y(v_i).
\]

As \( Y \) is a linear function of \( X \), \( F_r(q) \) (and \( F_b(q) \)) is of the form

\[
a_1 X - a_2 + X - a_3 + a_4,
\]

where \( X \) represents the \( x \)-coordinate of \( q \) which is bounded by \( x(v_k) \) and \( x(v_{k+1}) \), and \( a_1, a_2, a_3, a_4 \) are constants related to the local geometry.

Now the problem is to find the intersection of the two functions \( F_r() \) and \( F_b() \) at some optimal point \( q^*(e, i, j) \) (with respect to \( e \)). (Geometrically, \( q^*(e, i, j) \) is the point such that if one moves it by any small distance along \( e \), the maximum height of the two towers whose top must both see \( q^*(e, i, j) \) will be increased.) What we do is to first merge the red/blue interval points into a sorted list (which can be done in \( O(h + m) \) time), then we use prune and search to locate the interval which contains the intersection of \( F_r() \) and \( F_b() \). This can be done in \( O(h + m) = O(n) \) time: we first choose the \( \lfloor \frac{h + m}{2} \rfloor \)th interval point \( q' \), if \( F_r(q') > F_b(q') \) then we throw away all the interval points to the left of \( q' \); if \( F_r(q') < F_b(q') \) then we throw away all the interval points to the right of \( q' \); if \( F_r(q') = F_b(q') \) then we are done. This procedure continues until either we have \( F_r(q') = F_b(q') \) (in this case \( q' = q^*(e, i, j) \)) or we have only two sub-intervals left, one red and one blue. At this stage we simply solve an equation \( F_r(q) = F_b(q) \) to obtain \( q^*(e, i, j) \), which can be done in \( O(1) \) time as this is a 1-variable quadratic equation. To evaluate \( F_r(q') \) and \( F_b(q') \) in \( O(1) \) time, we just keep track of, for each interval point \( q' \) on \( e \), the previous and successive red and blue interval points. This can be easily done when we merge the red and blue interval points. Therefore, the time it takes to compute the optimal solution when both \( i, j \) and \( e \) are fixed is

\[
O(n) + O(n/2) + O(n/4) + \ldots + O(1) = O(n).
\]

It remains to show that the tops of the two towers can collectively cover \( T \). This can be done in a similar fashion as in case 1. We just need to compute the visibility region of (every edge of) \( T \) from \( u_i, u_j \) and then check whether each edge is collectively covered by \( u_i \) and \( u_j \). This can be done in \( O(n) \) time. If \( u_i, u_j \) cannot collectively cover \( T \), we simply throw them away as valid candidates. As we might have \( O(n) \) edges between \( v_i, v_j \) we have to try all of them and record all the pair of towers whose tops collectively cover \( T \). Among those recorded pairs we pick up the one whose height is the shortest as \( H_2(T, i, j) \). Therefore we have the following lemma.
Lemma 4. When $i, j$ are fixed, the optimal solution for the discrete 2-watchtower problem can be obtained in $O(n^2)$ time and $O(n)$ space.

With Lemma 4, it is straightforward to have the following theorem as we have $O(n^2)$ number of combinations of $i, j$.

**Theorem 1.** The optimal solution for the discrete 2-watchtower problem can be obtained in $O(n^3)$ time and $O(n)$ space.

Finally, we show that with parametric search [Me83] the discrete 2-watchtower problem can be solved in $O(n^3 \log^2 n)$ time. The height $h$ of two towers is a parameter of the decision problem which asks whether there exist a pair of vertices $v_i$ and $v_j$ of the terrain $T$ such that every point of the terrain is visible from towers of height $h$ located at $v_i$ and $v_j$. We need two algorithms, sequential and parallel, for the decision problem. Let $T_s$ be the running time of the sequential decision algorithm. (From the previous discussion, it is easy to see $T_s = O(n^3)$.) Let $T_p$ and $P$ denote the running time and the number of processors of the parallel algorithm respectively. The parametric searching scheme allows us to solve the optimization problem, i.e., the planar discrete 2-watchtower problem, in $O(PT_p + T_pT_s \log P)$ time. We use a parallel algorithm of Atallah et al. [ACW91] for computing visibility in a simple polygon. Their algorithm is based on a divide-and-conquer technique and can be implemented in the weak parallel model of computation of Valiant [Va75]. The running time is $O(\log n)$ using $O(n/\log n)$ processors. We just need a simplified version of the algorithm with $O(\log n)$ running time and $O(n)$ processors. We apply the algorithm in parallel for all pairs of $v_i, v_j$ using $O(n^3)$ processors. Hence $T_p = O(\log n)$ and $P = O(n^3)$.

**Theorem 2.** With parametric search, the optimal solution for the discrete 2-watchtower problem can be obtained in $O(n^3 \log^2 n)$ time.

Notice that so far it is not possible to actually implement an algorithm using parametric search. Therefore, this result is only of theoretical meaning.

### 2.2 Algorithm for the Planar 2-Watchtower Problem

We now discuss the solutions for the planar 2-watchtower problem. Our idea is basically the same as that for the discrete case. Again the following lemma, which presents a necessary condition for the planar two-watchtower problem, is easy to prove.

**Lemma 5.** If $\overline{uv}$ $(\overline{u'v'})$ achieves the optimal solution for the planar 2-watchtower problem then $u, u'$ collectively see $T$ and there exists a point $q$ on $S$ such that $h(q, v') \geq h(q, v)$ ($h(q, v) \geq h(q, v')$).

Similar to the discrete case, we also have two cases when $q$ is a vertex of $T$ and $q$ is on some edge of $T$. However, for both cases $v$ and $v'$ do not necessarily locate on a vertex of $T$ anymore. Let $L$ be the set of lines which are the extensions of edges of $T$. We compute the arrangement of $L$, $\mathcal{A}(L)$. We have the following lemma.
Lemma 6. In case 1, an optimal solution for the planar 2-watchtower \( \overline{uv} \) has the property that either \( v \) is a vertex of \( T \) or \( u \) is a vertex of \( A(L) \).

Proof. It suffices to prove the second part of the lemma. Assume that \( \overline{u^*v^*} \) is the optimal solution for the planar 2-watchtower problem of \( T \) such that \( v^* \) is not a vertex of \( T \) and \( u^* \) is not a vertex of \( A(L) \). We show below how to construct \( \overline{uv} \) such that every point of \( T \) which is covered only by \( u^* \) is also covered by \( u \) and \( |uv| \leq |u^*v^*| \).

By the definition of the problem, under case 1, \( u^* \) must be on a line \( l^* \) of \( L \) and \( v^* \) is on an edge \( e \) of \( T \) (Figure 3). If \( l^* \) and \( e \) are parallel then we can slide \( \overline{u^*v^*} \) to either a vertex of \( e \) or the closest vertex \( u^+ \) of \( A(L) \) on \( l^* \), which is defined by \( l^* \) and \( l^+ \) such that some points on \( T \) seen by \( u^+ \) are not seen by the top of the other watchtower. Let the resulting watchtower be \( \overline{uv} \), then \( |uv| = |u^*v^*| \) and either \( v \) is a vertex of \( T \) or \( u \) is a vertex of the arrangement \( A(L) \).

If \( l^* \) and \( e \) are not parallel then we can slide \( \overline{u^*v^*} \) (this time assume that we can always squeeze it) to either a vertex of \( e \) or the closest vertex \( u^+ \) of \( A(L) \) on \( l^* \), which is again defined by \( l^* \) and \( l^+ \) such that some points on \( T \) seen by \( u^+ \) are not seen by the top of the other watchtower and the height \( |u^*v^*| \) becomes shorter in the process of sliding. Let the resulting watchtower be \( \overline{uv} \), then \( |uv| < |u^*v^*| \) and either \( v \) is a vertex of \( T \) or \( u \) is a vertex of the arrangement \( A(L) \).

The proof of Lemma 6 naturally gives us an algorithm to solve the problem under case 1 in \( O(n^4) \) time, by converting the solutions of the \( O(n^2) \) discrete 2-watchtower problems, where the bases of the towers are on some vertices of \( T \), to the general problem. We thus have the following lemma.
Lemma 7. In case 1, the optimal solution for the planar 2-watchtower problem can be obtained in $O(n^4)$ time and $O(n^2)$ space.

Fig. 4. Handling case 2 for the general two-watchtower problem.

Case 2 is similar to the one we handle in the previous subsection. The difference is that now $v, v'$ can be on some edges of $T$. Let $v \in e', v' \in e'', q \in e$ and without loss of generality let $e' = v_i v_{i+1}, e'' = v_j v_{j+1}$ and $e = v_k v_{k+1}$ (Figure 4).

Now the function which defines the critical height $h(q, v), F_r(e', q)$, where $q \in e$ and $v \in e'$, is a piecewise decreasing function when $v$ is fixed. Let $q = (X_1, Y_1)$ and $u = (X_2, Y_2)$. $F_r(e, q)$ is of the form

$$a_1 (X_2 - a_2) \frac{X_1 - a_3}{X_1 - a_4} + a_5 X_2 + a_6, b_1 \leq X_1 \leq b_2, b_3 \leq X_2 \leq b_4, \ldots \ldots (A)$$

where $a_i (1 \leq i \leq 6), b_j (1 \leq j \leq 4)$ are constants related to local geometry. Similarly, the function $F_b(e'', q)$, where $q \in e$ and $v' \in e''$, defines the critical height $h(q, v')$. It is a piecewise increasing function when $v'$ is fixed and of the same form as (A). Notice that once $X_1$ is fixed we can easily maximize and minimize the value of (A) in $O(1)$ time. This evaluation will be used in the prune and search step. Similar to the previous section, the prune and search step stops when we have obtained an optimal interval point or we have located a red and a blue intervals such that $F_r(e, q), F_b(e, q)$ overlap (i.e., within the intersection of these two intervals the maximum of $F_r(e, q)$ is larger than the minimum of $F_b(e, q)$ and vice versa).

Finally, if we assume that the following equation with variables $X_1, X_2, X_3$,

$$a_1 (X_2 - a_2) \frac{X_1 - a_3}{X_1 - a_4} + a_5 X_2 + a_6 = a'_1 (X_3 - a'_2) \frac{X_1 - a'_3}{X_1 - a'_4} + a'_5 X_3 + a'_6, \ldots \ldots (B)$$
where \( a_i, a'_i (1 \leq i \leq 6) \) are constants and \( b_1 \leq X_1 \leq b_2, b_3 \leq X_2 \leq b_4, b'_3 \leq X_3 \leq b'_4 \) for constants \( b_1, b_2, b_3, b_4, b'_3, b'_4 \), can be solved in \( f_3 \) time (which can be done numerically in practice), then the merging of red/blue list, the subsequent prune and search step and the checking of whether \( u, u' \) can collectively cover \( T \) will all be done in \( O(n + f_3) \) time with \( e, e', e'' \) fixed. As we have \( O(n) \) candidates for \( e \) and \( O(n^2) \) number of combinations of \( e', e'' \), case 2 can be solved in \( O(n^4 + n^3 f_3) \) time. We hence have the following lemma.

**Lemma 8.** In case 2, if we assume that the above cubic equation (B) can be solved in \( f_3 \) time then the planar 2-watchtower problem can be solved in \( O(n^4 + n^3 f_3) \) time and \( O(n) \) space.

Therefore we have the following theorem.

**Theorem 3.** Assuming that the above cubic equation (B) can be solved in (approximately) \( f_3 \) time then the optimal solution for the planar 2-watchtower problem can be obtained in \( O(n^4 + n^3 f_3) \) time and \( O(n^2) \) space.

Finally, we show that with parametric search the planar 2-watchtower problem can be solved in \( O(n^4 \log^2 n) \) time. The height \( H \) of two towers is a parameter of the decision problem which asks whether there exist a pair of towers sliding on \( e' \) and \( e'' \) of the terrain \( T \) such that every point of the terrain is visible from towers of height \( H \) located on \( e' \) and \( e'' \). Let \( T_s \) be the running time of the sequential decision algorithm. It is easy to see that \( T_s = O(n^4) \): for each triple \( e, e', e'' \), there are \( O(n) \) event points (endpoints of the red/blue intervals) and checking whether the terrain \( T \) is completely visible can be done in \( O(n) \) time. Let \( T_p \) and \( P \) denote the running time and the number of processors of the parallel algorithm respectively. It is easy to see that \( T_p = O(\log n) \) and \( P = O(n^4) \): we use a processor for each \( e, e', e'' \) and an event point on \( e \) (this uses a total of \( O(n^4) \) processors); moreover, with these processors it is easy to check in parallel whether \( T \) can be completely covered in \( O(\log n) \) time. Hence the total running time is \( O(n^4 \log^2 n) \). (Again, this result is only of theoretical meaning.)

**Theorem 4.** With parametric search, the optimal solution for the planar 2-watchtower problem can be obtained in \( O(n^4 \log^2 n) \) time.

### 3 Closing Remarks

In this paper, we present several polynomial time algorithms to solve the planar 2-watchtower problems. An immediate open problem is how to solve the 2-watchtower problem in 3D efficiently. (We remark that for the discrete 2-watchtower problem in 3D, where the towers must be located at the vertices of the terrain, it is possible to use a necessary condition similar to Lemma 2 to obtain an \( O(n^6) \) time solution.)
References


Efficient Generation of Triconnected Plane Triangulations

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Abstract. A “rooted” plane triangulation is a plane triangulation with one designated vertex on the outer face. In this paper we give a simple algorithm to generate all triconnected rooted plane triangulations with at most \( n \) vertices. The algorithm uses \( O(n) \) space and generates such triangulations in \( O(1) \) time per triangulation without duplications. The algorithm does not output entire triangulations but the difference from the previous triangulation. By modifying the algorithm we can generate all triconnected rooted plane triangulations having exactly \( n \) vertices including exactly \( r \) vertices on the outer face in \( O(r^2) \) time per triangulation, while the previous best algorithm generates such triangulations in \( O(n^2) \) time per triangulation. Also we can generate without duplications all triconnected (non-rooted) plane triangulations having exactly \( n \) vertices including exactly \( r \) vertices on the outer face in \( O(r^2n) \) time per triangulation, and all maximal planar graphs in \( O(n^3) \) time per graph.

1 Introduction

Generating all graphs with some property without duplications has many applications, including unbiased statistical analysis [M98]. A lot of algorithms to solve these problems are already known [A96, B80, LN01, M98, W86, etc]. See nice textbooks [G93, KS98]. In this paper we wish to generate all triconnected “rooted” plane triangulations, which will be defined precisely in Section 2, with at most \( n \) vertices. It is known that every triconnected planar graph has a unique embedding on a sphere only up to mirror copy [HW74].

To solve these all-graph-generating problems some types of algorithms are known. Classical method algorithms [G93, p57] first generate all the graphs with given property allowing duplications, but output only if the graph has not been output yet. Thus this method requires quite a huge space to store a list of graphs that have already been output. Furthermore, checking whether each graph has already been output requires a lot of time. Orderly method algorithms [G93, p57] need not to store the list, since they output a graph only if it is a “canonical” representative of each isomorphism class. Reverse search method algorithms [A96] also need not to store the list. The idea is to implicitly define a connected graph \( H \) such that the vertices of \( H \) correspond to the graphs with
the given property, and the edges of $H$ correspond to some relation between the graphs. By traversing an implicitly defined spanning tree of $H$, one can find all the vertices of $H$, which correspond to all the graphs with the given property.

The main idea of our algorithm is that for some problems we can define a tree (not a general graph) as the graph $H$ of reverse search method. Thus our algorithm does not need to find a spanning tree of $H$, since $H$ itself is a tree. With some other ideas we give the following two simple but efficient algorithms.

Our first algorithm generates all triconnected rooted plane triangulations with at most $n$ vertices. A rooted plane triangulation means a plane triangulation with one designated “root” vertex on the outer face. For instance there are four triconnected rooted plane triangulations with at most five vertices including exactly three vertices on the outer face, as shown in Fig. 1(a). The root vertices are depicted by white circles. However, there are only two triconnected plane triangulations with at most five vertices including exactly three vertices on the outer face. See Fig. 1(b). The algorithm uses $O(n)$ space and runs in $O(f(n))$ time, where $f(n)$ is the number of nonisomorphic triconnected rooted plane triangulations with at most $n$ vertices. The algorithm generates triangulations without duplications. So the algorithm generates each triangulation in $O(1)$ time on average. The algorithm does not output entire triangulations but the difference from the previous triangulation.

By modifying our first algorithm we can generate without duplications all triconnected rooted plane triangulations having exactly $n$ vertices including exactly $r$ vertices on the outer face. The algorithm uses $O(n)$ space and runs in $O(r \cdot f(n, r))$ time, where $f(n, r)$ is the number of nonisomorphic such triangulations. So the algorithm generates each triangulation in $O(r)$ time on average, while the previous best algorithm [A96] generates such triangulations in $O(n^2)$ time per triangulation.

![Fig. 1.](image)

Fig. 1. (a) Triconnected rooted plane triangulations, and (b) triconnected plane triangulations.

Also we can generate all triconnected (non-rooted) plane triangulations having exactly $n$ vertices including exactly $r$ vertices on the outer face in $O(r^2 n)$ time (on average) per triangulation, and all maximal planar graphs in $O(n^3)$ time per graph. Our algorithm is simple and does not need the complicated theoretical linear-time plane-graph-isomorphism algorithm in [HW74].

For full version of the paper see [N01]. We also solved similar problems for biconnected cases [LN01].
The rest of the paper is organized as follows. Section 2 gives some definitions. Section 3 shows a tree structure among triconnected rooted plane triangulations. Section 4 presents our first algorithm. By modifying the algorithm we give three more algorithms in Section 5.

2 Preliminaries

Let $G$ be a connected graph with $n$ vertices. An edge connecting vertices $x$ and $y$ is denoted by $(x, y)$. The degree of a vertex $v$, denoted by $d(v)$, is the number of neighbors of $v$ in $G$. A cut is a set of vertices whose removal results in a disconnected graph or a single-vertex graph $K_1$. The connectivity $\kappa(G)$ of a graph $G$ is the cardinality of the minimum number of vertices consisting a cut. $G$ is $k$-connected if $k \leq \kappa(G)$.

A graph is planar if it can be embedded in the plane so that no two edges intersect geometrically except at a vertex to which they are both incident. A plane graph is a planar graph with a fixed planar embedding. A plane graph divides the plane into connected regions called faces. The unbounded face is called the outer face, and other faces are called inner faces. We regard the contour of a face as the clockwise cycle formed by the vertices and edges on the boundary of the face. We denote the contour of the outer face of plane graph $G$ by $C_o(G)$. A vertex is an outer vertex of $G$ if it is on $C_o(G)$, and an inner vertex otherwise. An edge is an outer edge of $G$ if it is on $C_o(G)$, and an inner edge otherwise. A plane graph is called a plane triangulation if each inner face has exactly three edges on its contour. A rooted plane triangulation is a plane triangulation with one designated vertex on $C_o(G)$. The designated vertex is called the root vertex.

Let $G$ be a triconnected plane triangulation. Assume $G$ has a cut $S$ consisting exactly three vertices, say $x, y, z$. Let $G'$ be the graph derived from $G$ by removing $S$. Let $W$ be the vertex set of a connected component of $G'$. Let $G''$ be the plane subgraph of $G$ induced by $S \cup W$. Since $G$ is triconnected each vertex in $S$ has at least one neighbor in $W$. Since $G$ is a plane triangulation, all vertices in $S$ are on the same face, say $F$, of $G''$. We now have the following lemma.

Lemma 1. Let $P$ be the path on the contour of $F$ between $x$ and $y$ not containing $z$. Then the followings hold. (a) If $x$ and $y$ are inner vertices of $G$, then $G$ has an (inner) edge $(x, y)$. (b) If $x$ is an inner vertex and $y$ is an outer vertex of $G$, then $G$ has an (inner) edge $(x, y)$. (c) If $x$ and $y$ are outer vertices of $G$, then all vertices on $P$ are also outer vertices.

Thus if $G \neq K_4$, then at most two vertices of $S$ are outer vertices, because by (c) above if all three vertices in $S$ are outer vertices, then $S$ cannot be a cut. If $S$ has at most one outer vertex, then a cut $S = \{x, y, z\}$ of $G$ is called a separating triangle, and by (a) and (b) above $G$ has three edges $(x, y)$, $(y, z)$ and $(z, x)$. If $S$ has exactly two outer vertex, say $x$ and $y$, and exactly one inner vertex, say $z$, then a cut $S = \{x, y, z\}$ of $G$ is called a separating fan with center $z$, and by (b) above $G$ has two edges $(x, z)$, $(y, z)$. 
Then we define an operation on a plane triangulation. Assume \((u,v)\) is an inner edge of a plane triangulation \(G\). Let \(u,v,x\) and \(v,u,y\) be the two faces having \((u,v)\) on its contour. By (i) deleting the edge \((u,v)\), (ii) merging \(u\) and \(v\) into a new vertex \(w\), and (iii) replacing edges \((u,x)\) and \((v,x)\) by \((w,x)\), and edges \((u,y)\) and \((v,y)\) by \((w,y)\), as shown in Fig. 2(a), we can have a new plane graph. This operation is called a contraction of inner edge \((u,v)\).

Similarly, for an outer edge \((u,v)\) of a plane triangulation \(G\), let \(u,v,x\) be the face having \((u,v)\) on its contour. By (i) deleting the edge \((u,v)\), (ii) merging \(u\) and \(v\) into a new vertex \(w\), and (iii) replacing edges \((u,x)\) and \((v,x)\) by \((w,x)\), we can have a new plane graph. This operation is called a contraction of outer edge \((u,v)\).

Let \(G\) be a triconnected plane triangulation. If \(G\) has a separating triangle \(\{u,v,z\}\), then the contraction of \((u,v)\) produces “parallel” edge \((w,z)\), so the resulting plane graph is not a plane triangulation anymore. If \(G\) has a separating fan \(\{u,v,z\}\) with center \(v\), then after contracting \((u,v)\), the resulting plane graph has a cut \(\{w,z\}\), and is not triconnected anymore. If \(G\) has no cut consisting of \(u,v\) and exactly one more vertex, then the contraction of \((u,v)\) produces again a triconnected plane triangulation. We say such edge \((u,v)\) is contractible.

![Fig. 2. (a) A contraction and (b) a contraction sequence.](image)

3 The Contracting Sequence and the Genealogical Tree

Let \(S_n\) be the set of all triconnected rooted plane triangulations with at most \(n\) vertices. Here we explain a tree structure among the triangulations in \(S_n\).

Let \(G\) be a triangulation in \(S_n\) but not \(K_4\). Let \(C_0(G) = v_1, v_2, \ldots, v_k\), and \(v_1\) be the root vertex of \(G\). A contractible edge \((v_1,w_i)\) is the first contractible edge of \(G\) if none of \((v_1,w_2),(v_1,w_3),\ldots,(v_1,w_{i-1})\) is contractible edge, where \((v_1,v_k) = (v_1,w_1),(v_1,w_2),\ldots,(v_1,w_{d(v_1)}) = (v_1,v_2)\) are the edges incident to \(v_1\), and assume that they appear around \(v_1\) counterclockwise in this order.

**Lemma 2.** Every triconnected rooted plane triangulation except \(K_4\) has the first contractible edge.

For each triconnected rooted plane triangulation \(G\) in \(S_n\) except \(K_4\), contracting the first contractible edge from \(G\) results in also a triconnected rooted plane triangulation, denoted by \(P(G)\), in \(S_n\) with one less vertices. Thus we can define the triangulation \(P(G)\) in \(S_n\) for each \(G\) in \(S_n\) except \(K_4\). We say \(G\) is a child triangulation of \(P(G)\).
Given a triangulation $G$ in $S_n$, by repeatedly contracting the first contractible edge, we can have the unique sequence $G, P(G), P(P(G)), \cdots$ of triangulations in $S_n$ which eventually ends with $K_4$. See an example in Fig. 2(b), in which the first contractible edges are depicted by thick lines.

By merging those sequences we can have the genealogical tree $T_n$ of $S_n$ such that the vertices of $T_n$ correspond to the triangulations in $S_n$, and each edge corresponds to each relation between some $G$ and $P(G)$. For instance $T_6$ is shown in Fig. 3 in which the first contractible edges are depicted by thick lines. We call the vertex in $T_n$ corresponding to $K_4$ the root of $T_n$.

4 Algorithm

Given $S_n$ we can construct $T_n$ by the definition, possibly with huge space and much running time. However, how can we construct $T_n$ efficiently only given an integer $n$? Our idea is by reversing the contracting procedure as follows.

Given a triconnected rooted plane triangulation $G$ in $S_n$ with at most $n-1$ vertices, we wish to find all child triangulations of $G$. Let $C_0(G) = v_1, v_2, \cdots, v_k$, and $v_1$ be the root vertex of $G$. Let $(v_1, v_k) = (v_1, w_1), (v_1, w_2), \cdots, (v_1, w_d(v_1)) = (v_1, v_2)$ be the edges incident to $v_1$, and assume that they appear around $v_1$ counterclockwise in this order. Assume $(v_1, w_s)$ be the first contractible edge of $G$. Since especially only $K_4$ has no contractible edge, for convenience, we regard $(v_1, w_s) = (v_1, v_2)$ as the first contractible edge for $K_4$.

We denote by $G_i(i, j)$, where $1 \leq i < j \leq d(v_1)$, the rooted plane triangulation obtained from $G$ by (1) removing edges $(v_1, w_{i+1}), (v_1, w_{i+2}), \cdots, (v_1, w_{j-1})$
(thus if \( i + 1 = j \) we need not remove any edge), (2) adding new vertex \( v \) on the inner face of \( G \) containing \( v_1, w_i, w_j \), (3) adding an edge \((v_1, v)\) in the face, and (4) adding \( j - i + 1 \geq 2 \) edges \((v, w_i), (v, w_{i+1}), \ldots, (v, w_j)\), as shown in Fig. 4(a). \( G_i(i,j) \), where \( 1 \leq i < j \leq \ell(v_1) \), is a child triangulation of \( G \) if and only if \((v_1, v)\) is the first contractible edge of \( G_i(i,j) \). (This corresponds to a contraction of an inner edge. See Fig. 3)

Also, we denote by \( G_o(i,j) \), where \( i = \ell(v_1) \) and \( 3 \leq j \leq \ell(v_1) \), the rooted plane triangulation obtained from \( G \) by (1) removing edges \((v_1, w_j), (v_1, w_{j+1}), \ldots, (v_1, w_{\ell(v_1)})\), (2) adding new vertex \( v \) on the outer face of \( G \), (3) adding edges \((v_1, v)\) and \((v, v_2)\), and (4) adding edges \((v, w_{j-1}), (v, w_j), \ldots, (v, w_{\ell(v_1)})\), as shown in Fig. 4(b). \( G_o(i,j) \), where \( i = \ell(v_1) \) and \( 3 \leq j \leq \ell(v_1) \), is a child triangulation of \( G \) if and only if \((v_1, v)\) is the first contractible edge of \( G_o(i,j) \). (This corresponds to a contraction of an outer edge. See Fig. 3)

![Fig. 4. Illustration for (a) \( G_i(i,j) \) and (b) \( G_o(i,j) \).](image)

Thus \( G_i(i,j) \) is a child triangulation of \( G \) having an inner edge as the first contractible edge, while \( G_o(i,j) \) is a child triangulation of \( G \) having an outer edge as the first contractible edge. And each child triangulation is either \( G_i(i,j) \) or \( G_o(i,j) \) for some \( i \) and \( j \).

For each \( t \), \( 2 \leq t \leq \ell(v_1) - 1 \), we denote by \( q(t) \) the largest index such that \((w_t, w_{q(t)})\) is an edge (chord) of \( G \) if such edge exists, and \( q(t) = t \) otherwise. Now we check whether each \( G_i(i,j) \) is a child triangulation of \( G \), and check also for each \( G_o(i,j) \), as follows. Let \((v_1, w_s)\) be the first contractible edge of \( G \).

**Case 1:** \( G_i(i,j) \), where \( 1 \leq i < j \leq \ell(v_1) \). We have four subcases.

**Case 1(a):** \( j \leq s \).

In this case \((v_1, v)\) is the first contractible edge of \( G_i(i,j) \), and \( P(G_i(i,j)) = G \).

**Case 1(b):** \( i < s < j \).

If \( G \) has a separating fan containing \( v_1 \) and \( w_i \), then \((v_1, v)\) is the first contractible edge for each \( j, s < j \leq \ell(v_1) \) and \( P(G_i(i,j)) = G \). Otherwise, if \( j \leq q(i) \) then \((v_1, v)\) is the first contractible edge and \( P(G_i(i,j)) = G \). Otherwise \((v_1, w_i)\) is a contractible edge of \( G(i,j) \), and so \( P(G_i(i,j)) \neq G \).

**Case 1(c):** \( i = s \).

If \( j = i+1 \), then \((v_1, v)\) is the first contractible edge of \( G_i(i,j) \) (since \( v_1, w_s, w_{s+1} \) become a separating triangle) and \( P(G_i(i,j)) = G \). Otherwise, \( j \geq i + 2 \) holds then \((v_1, w_s)\) is still contractible in \( G_i(i,j) \), and \( P(G_i(i,j)) \neq G \).

**Case 1(d):** \( i > s \).

\((v_1, v)\) is not the first contractible edge of \( G_i(i,j) \). Thus \( P(G_i(i,j)) \neq G \).
**Case 2:** \( G_o(i,j) \), where \( i = d(v_1) \) and \( 3 \leq j \leq d(v_i) \). We have three subcases.

**Case 2(a):** \( s = d(v_1) \).

(Now each \( w_2, w_3, \ldots, w_{d(v_1)} \) is contained in some separating fan.) For each \( G_o(i,j), 3 \leq j \leq d(v_1), (v_1,v) \) is the first contractible edge, and \( P(G_o(i,j)) = G \).

**Case 2(b):** \( s = d(v_1) - 1 \).

For each \( G(i,j), 3 \leq j \leq d(v_1), (v_1,v) \) is the first contractible edge, and \( P(G_o(i,j)) = G \). Note that in \( G_o(i,d(v_1)), v_1, w_{d(v_1)} - 1, v_2 \) become a separating fan, and other cases are trivial.

**Case 2(c):** \( s < d(v_1) - 1 \).

If \( j - 1 \geq s \) then \( (v_1, w_s) \) is still contractible in \( G_o(i,j) \), thus \( (v_1,v) \) is not the first contractible edge and \( P(G_o(i,j)) \neq G \). Otherwise if \( G \) has no separating fan containing \( v_1 \) and \( w_{j-1} \) then \( (v_1, w_{j-1}) \) is contractible in \( G_o(i,j) \), thus \( (v_1,v) \) is not the first contractible edge, and \( P(G_o(i,j)) \neq G \). Otherwise, for each \( j, 3 \leq j \leq s, (v_1,v) \) is the first contractible edge in \( G_o(i,j) \), and \( P(G_o(i,j)) = G \).

Based on the case analysis above we can find all child triangulations of given triangulation in \( S_n \). If \( G \) has \( k \) child triangulations, then we can find them in \( O(k) \) time. This is an intuitive reason why our algorithm generates triangulations in \( O(1) \) time per triangulation.

And recursively repeating this process from the root of \( T_n \) corresponding to \( K_4 \) we can traverse \( T_n \) without constructing whole \( T_n \). During the traversal of \( T_n \), we assign a label \((i,j)\) to each edge connecting \( G \) and either \( G_i(i,j) \) or \( G_o(i,j) \) in \( T_n \), as shown in Fig. 3. Each label denotes how to generate a child triangulation of \( G \), and each sequence of labels on a path starting from the root specifies a triangulation in \( S_n \). For instance \((1, 2), (1, 2)\) specify the uppermost triangulation in Fig. 3. During our algorithm we will maintain these labels only on the path from the root to the “current” vertex, because those are enough information to generate the “current” triangulation. To generate next triangulation, we need to maintain some more information (the first contractible edge \((v_1, w_s)\), and \( w_q(t) \) for each \( 1 < t < s \), etc.) only for the triangulations on the “current” path, which has length at most \( n \). This is an intuitive reason why our algorithm uses only \( O(n) \) space, while the number of triangulations may not be bounded by a polynomial in \( n \).

**Procedure find-all-child-triangulations**

begin
1. output \( G \) { Output the difference from the previous triangulation}
2. if \( G \) has exactly \( n \) vertices then **return**
3. for \( i = 1 \) to \( s - 1 \)
4. for \( j = i + 1 \) to \( s \)
5. find-all-child-triangulations\((G_i(i,j))\) { Case 1(a)}
6. for \( i = 1 \) to \( s - 1 \)
7. if \( G \) has a separating fan containing \( v_1 \) and \( w_i \) then
8. for \( j = s + 1 \) to \( d(v_1) \)
9. find-all-child-triangulations\((G_i(i,j))\) { Case 1(b)}
10. else then
for $j = s + 1$ to $q(i)$
  find-all-child-triangulations\($G_i(i, j)$\) \{ Case 1(b)\}
find-all-child-triangulations\($G_i(s, s + 1)$\) \{ Case 1(c)\}
if $s = d(v_1)$ or $d(v_1) - 1$ then
  for $j = 3$ to $d(v_1)$
    find-all-child-triangulations\($G_o(d(v_1), j)$\) \{ Case 2(a) and (b)\}
else then
  for $j = 3$ to $s$
    if $G$ has a separating fan containing $v_1$ and $w_{j-1}$ then
      find-all-child-triangulations\($G_o(d(v_1), j)$\) \{ Case 2(c)\}
end

Algorithm find-all-triangulations\($n$\)
begin
  output $K_4$
  if $n = 4$ then return
  $G = K_4$
  find-all-child-triangulations\($G_i(1, 2)$\)
  find-all-child-triangulations\($G_i(1, 3)$\)
  find-all-child-triangulations\($G_i(2, 3)$\)
  find-all-child-triangulations\($G_o(3, 3)$\)
end

Theorem 1. The algorithm uses $O(n)$ space and runs in $O(f(n))$ time, where $f(n)$ is the number of nonisomorphic triconnected rooted plane triangulations with at most $n$ vertices.

5 Modification of the Algorithm

Then we consider our second problem. Let $S_{\equiv r}^n$ be the set of triconnected rooted plane triangulations having exactly $n$ vertices including exactly $r$ outer vertices. We wish to generate all triangulations in $S_{\equiv r}^n$ without duplications. Clearly all such triangulations are in $T_n$ but with other triangulations. How can we output only triangulations in $S_{\equiv r}^n$, furthermore efficiently? Our idea is by pruning $T_n$, in the following way.

For each triangulation $G$ in $S_n$ but $K_4$, if the first contractible edge is an inner edge then $P(G)$ is a triangulation in $S_n$ having the same number of outer vertices and one less inner vertices, otherwise the first contractible edge is an outer edge and $P(G)$ is a triangulation in $S_n$ having one less outer vertices and the same number of inner vertices. Thus if $G$ in $S_n$ has exactly $n'$ vertices including exactly $r'$ outer vertices, then every “descendant” triangulation of $G$ in $T_n$ can have at most $r' + (n - n')$ outer vertices. Thus if $r > r' + (n - n')$ then none of descendant triangulations of $G$ is in $S_{\equiv r}^n$. Also if $G$ has $r' > r$ outer vertices then every descendant triangulation of $G$ has $r'$ or more outer vertices, and so never in $S_{\equiv r}^n$. We prune such triangulations from $T_n$. 
We prune $T_n$ further based on the following lemma. We need some definitions here. Let $S^i_n$ be the set of triconnected rooted plane triangulations with $n' < n$ vertices including exactly $r' = r$ outer vertices (so it already has enough outer vertices), and $S^o_n$ be the set of triconnected rooted plane triangulations with $n' < n$ vertices including exactly $r' = r - (n - n')$ outer vertices (so it already has enough inner vertices). Let $S^r_n$ be the set of triconnected rooted plane triangulations with $n' < n$ vertices including $r'$, $r - (n - n') < r' < r$ outer vertices. Let $G$ be a triangulation in $S_n$, $C_0(G) = v_1, v_2, \ldots, v_k$, and $v_1$ be the root vertex of $G$, $(v_1, v_k) = (v_1, w_1), (v_1, w_2), \ldots, (v_1, w_d(v_1)) = (v_1, v_2)$ be the edges incident to $v_1$, and assume that they appear around $v_1$ counterclockwise in this order. We call $w_2$ the pivot of $G$. If $G$ has a separating fan containing both the root and the pivot of $G$, then it is called the rp-fan of $G$.

**Lemma 3.** Let $G$ be a triangulation in $S_n$ with $n' < n$ vertices including $r'$ outer vertices.

(a) If $G$ has the rp-fan, then (a1) if $G$ has a child triangulation $G_i(1, j)$ for some $j$, then $G_i(1, j)$ has no rp-fan, (a2) if $G$ has a child triangulation $G_i(2, j)$ for some $j$, then $G_i(2, j)$ has the rp-fan, (a3) if $G$ has a child triangulation $G_i(1, j)$ for some $i \geq 3$ and $j$, then $G_i(1, j)$ has the rp-fan, and (a4) if $G$ has a child triangulation $G_o(i, j)$ for some $i$ and $j$, then $G_o(i, j)$ has the rp-fan. 

(b) If $G$ has no rp-fan, then (b1) if $G$ has a child triangulation $G_i(1, j)$ for some $j$, then $G_i(1, j)$ has no rp-fan, (b2) $G$ has no child triangulation $G_i(i, j)$ for any $i \geq 2$ and $j$, and (b3) $G$ has no child triangulation $G_o(i, j)$ for any $i$ and $j$.

**Lemma 4.** Let $G$ be a triangulation in $S_n$ with $n' < n$ vertices including $r', r - (n - n') \leq r' < r$ outer vertices. If $G$ has no rp-fan then $G$ has no descendant triangulation in $S^r_{=n'}$. If $G$ has the rp-fan then $G$ has at least one descendant triangulation in $S^r_{=r}$.

**Proof.** By repeatedly applying Lemma 3(b), if $G$ has no rp-fan then each descendant triangulation of $G$ has no rp-fan and having exactly $r' < r$ outer vertices. Thus none of them are in $S^r_{=n'}$. While if $G$ has the rp-fan then the first contractible edge is not $(v_1, w_2)$ and then each $G_o(d(v_1), j)$ is a child triangulation of $G$ if and only if, for each $w = w_2, w_3, \ldots, w_{j-1}$, $G$ has a separating fan containing $w$ and $v_1$. Such child triangulation is always exist for $j = 3$. \hfill $\square$

Now we define a new tree, called the genealogical tree $T^r_n$. We say a vertex $v$ in a tree with root $r$ has depth $d$ if the path connecting $v$ and $r$ on the tree has exactly $d$ edges. Define $T^r_n$ as the tree such that (i) the vertices of $T^r_n$ having depth $n - 4$ correspond to the triangulations in $S^r_{=n'}$, (ii) the vertices of $T^r_n$ having depth at most $n - 5$ correspond to the triangulations in $S^r_n \cup S^i_n \cup S^o_n$, and (iii) each edge corresponds to each relation between some $G$ and $P(G)$. The vertex in $T^r_n$ corresponding to $K_4$ is called the root of $T^r_n$. Given a triangulation $G$ in $S^r_n \cup S^i_n \cup S^o_n$, we can find all child triangulations of $G$ efficiently, by a similar case analysis to the one in the previous section. If $G$ has $k$ child triangulations, then again we can find them in $O(k)$ time.
Procedure find-all-child-triangulations2\((G)\)
begin
1 Assume \(G\) has \(n'\) vertices including \(r'\) outer vertices.
2 if \(G\) has exactly \(n\) vertices then output \(G\) return
3 if \(r' < r\) and \(r' + (n - n') > r\) then
   { We need more both inner and outer vertices. }
4 for each \(G_i(i, j)\) with \(i \geq 2\) satisfying (a2) or (a3)
5 find-all-child-triangulations2\((G_i(i, j))\)
6 for each \(G_o(i, j)\) satisfying (a4)
7 find-all-child-triangulations2\((G_o(i, j))\)
8 else if \(r' < r\) and \(r' + (n - n') = r\) then
   { We need to increase only outer vertices. }
9 for each \(G_o(i, j)\) satisfying (a4)
10 find-all-child-triangulations2\((G_o(i, j))\)
11 else \(r' = r\) holds
12 for each \(G_i(i, j)\) satisfying (a1), (a2), (a3) or (b1)
13 find-all-child-triangulations2\((G_i(i, j))\)
end

Lemma 5. Let \(f(n, r)\) be the number of triangulations in \(S_{=}^{=n} r\). Then \(T_n^r\) has at most \(r \cdot f(n, r)\) vertices.

Proof. (Sketch) Each vertex of \(T_n^r\) corresponds to a triangulation in \(S_n^r \cup S_n^o \cup S_o^r \cup S_o^r\). Each vertex with depth \(n - 4\) in \(T_n^r\) corresponds to a triangulation in \(S_{=}^{=n} r\). Each “ancestor” triangulation \(G\) of a triangulation \(G' \in S_{=}^{=n} r\) has two or more child triangulation if \(G \in S_n^r \cup S_n^o\) (since \(G_i(1, 2)\) and \(G_i(1, 3)\) are always child triangulation of \(G\)), and has one or more child triangulation if \(G \in S_n^r\) and \(G\) has the rp-fan (since \(G_i(d(v_1), 3)\) is always child triangulation of \(G\)). And each \(G \in S_o^r\) has at most \(r - 4\) “ancestor triangulations which are in \(G \in S_o^r\). Thus \(T_n^r\) has at most \(r \cdot f(n, r)\) vertices. □

Since the algorithm needs only a constant time of computation for each edge of \(T_n^r\), we have the following lemma.

Lemma 6. The algorithm uses \(O(n)\) space and runs in \(O(r \cdot f(n, r))\) time, where \(f(n, r)\) is the number of nonisomorphic triconnected rooted plane triangulations having exactly \(n\) vertices including exactly \(r\) vertices on the outer face.

We modify our second algorithm further so that it output all triconnected (non-rooted) plane triangulations having exactly \(n\) vertices including exactly \(r\) vertices on the outer face, as follows.

At each leaf \(v\) of the genealogical tree \(T_n^r\), the triangulation \(G\) corresponding to \(v\) is checked whether the contracting sequence of \(G\) with the root vertex is the lexicographically first one among the \(r\) contracting sequences of \(G\) for \(r\) choices of the root vertex on \(C_o(G)\), and only if so \(G\) is output. Thus we can output only canonical representative of each isomorphism class.
The algorithm uses $O(n)$ space and runs in $O(r^2n \cdot g(n, r))$ time, where $g(n, r)$ is the number of nonisomorphic triconnected \((\text{non-rooted})\) plane triangulations having exactly $n$ vertices including exactly $r$ vertices on the outer face.

A triconnected (non-rooted) plane triangulations $G$ is maximal if $G$ has exactly 3 vertices on $C_o(G)$. For every maximal planar graph by choosing the outer face and the root vertex, there are exactly $2m = 3n - 6$ (triconnected) rooted plane triangulation $G$ with exactly 3 vertices on the outer face, where $m$ is the number of edges of $G$. We modify the algorithm further as follows. At each leaf $v$ of the genealogical tree $T^3_n$, the triangulation $G$ corresponding to $v$ is checked whether the contracting sequence of $G$ with the rooted vertex is the lexicographically first one among the $2m$ contracting sequences of $G$ (3 choice of the root vertex on $C_o(G)$ for each of $2m/3$ choice of the outer face of $G$), and only if so $G$ is output.

Theorem 2. The modified algorithm generate all maximal planar graphs in $O(n^3 \cdot h(n))$ time, where $h(n)$ is the number of nonisomorphic maximal planar graphs with exactly $n$ vertices. The algorithm uses $O(n)$ space.

References

Packing Two Disks
into a Polygonal Environment*

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Abstract. We consider the following problem. Given a polygon $P$, possibly with holes, and having $n$ vertices, compute a pair of equal radius disks that do not intersect each other, are contained in $P$, and whose radius is maximized. Our main result is a simple randomized algorithm whose expected running time, on worst case input, is $O(n \log n)$. This is optimal in the algebraic decision tree model of computation.

1 Introduction

Let $P$ be a polygon, possibly with holes, and having $n$ vertices. We consider the following problem, which we call 2-DISK: Find a pair of disks with radius $r^*$ that do not intersect each other, are contained in $P$, and such that $r^*$ is maximized. Biedl et al. [5] gave an $O(n^2)$ time algorithm to solve this problem.

Special cases of 2-DISK have been studied previously. When $P$ is a convex polygon, Bose et al. [6] describe a linear time algorithm and Kim and Shin [10] describe an $O(n \log n)$ time algorithm. For simple polygons (i.e. polygons without holes), Bespamyatnikh [4] gives an $O(n \log^3 n)$ time algorithm based on the parametric search paradigm [11].

Another special case occurs when the holes of $P$ degenerate to points. This is known as the maximin 2-site facility location problem [3,9]. In this formulation we can think of the centers of the two disks as obnoxious facilities such as smokestacks, or nuclear power plants, and the points as population centers. The goal is maximize the distance between each facility and the nearest population center. Katz et al. [9] give an $O(n \log n)$ time algorithm for the decision version of the 2-site facility location problem in which one is given a distance $d$ and asked if there exists a placement of 2 non-intersecting disks of radius $d$, each contained in $P$ such that no point is included in either of the disks.

In this paper we present a simple randomized algorithm for the general case in which $P$ is not necessarily convex and may contain holes. Our algorithm runs in $O(n \log n)$ expected time. It can also be used to solve the optimization version of the 2-site maximin facility location problem in $O(n \log n)$ time. Finally we

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observe that, when we allow polygons with holes, \( \Omega(n \log n) \) is a lower bound for 2-disk by a simple reduction from MAX-GAP.

The remainder of the paper is organized as follows: Section 2 reviews definitions and previous results regarding the medial-axis. Section 3 describes our algorithm. Section 4 summarizes and concludes with an open problem.

2 The Medial-Axis

For the remainder of this paper, \( P \) will be a polygon, possibly with holes, and having \( n \) vertices. The medial-axis \( M(P) \) of \( P \) is the locus of all points \( p \) for which there exists a disk centered at \( p \), contained in \( P \), and which intersects the boundary of \( P \) in two or more points. See Fig. 1 for an example. Alternatively, \( M(P) \) is a portion of the Voronoi diagram of the open line segments and vertices defined by the edges of \( P \). To be more precise, we need to remove the Voronoi edges that are outside \( P \) and those associated with an edge and one of its endpoints. It is well known that the medial-axis consists of \( O(n) \) straight line segments and parabolic arcs.

Algorithmically, the medial-axis is well understood. There exists an \( O(n) \) time algorithm [7] for computing the medial-axis of a polygon without holes and \( O(n \log n) \) time algorithms for computing the medial-axis of a polygon with holes [2]. Furthermore, these algorithms can compute a representation in which each segment or arc is represented as a segment or arc in \( \mathbb{R}^3 \), where the third dimension gives the radius of the disk that touches two or more points on the boundary of \( P \).

We say that a point \( p \in P \) supports a disk of radius \( r \) if the disk of radius \( r \) centered at \( p \) is contained in \( P \). We call a vertex, parabolic arc or line segment \( x \) of \( M(P) \) an elementary object if the radius of the largest disk supported by \( p \in x \) is monotone as \( p \) moves from one endpoint of \( x \) to the other. Each edge of \( M(P) \) can be split into two elementary objects. Thus, \( M(P) \) can be split into \( O(n) \) elementary objects whose union is \( M(P) \).
3 The Algorithm

Next we describe a randomized algorithm for 2-disk with \( O(n \log n) \) expected running time. We begin by restating 2-disk as a problem of computing the diameter of a set of elementary objects under a rather unusual distance function. We then use an algorithm based on the work of Clarkson and Shor [8] to solve this problem in the stated time.

The following lemma, of which similar versions appear in Bose et al. [6] and Biedl et al. [5], tells us that we can restrict our search to disks whose centers lie on \( M(P) \).

**Lemma 1.** Let \( D_1 \) and \( D_2 \) be a solution to 2-disk which maximizes the distance between \( D_1 \) and \( D_2 \) and let \( p_1 \) and \( p_2 \) be the centers of \( D_1 \) and \( D_2 \), respectively. Then \( D_1 \) and \( D_2 \) each intersect the boundary of \( P \) in at least two points and hence \( p_1 \) and \( p_2 \) are points of \( M(P) \).

**Proof.** Refer to Fig. 2. Suppose that one of the disks, say \( D_1 \), intersects the boundary of \( P \) in at most one point. Let \( o_1 \) be this point, or if \( D_1 \) does not intersect the boundary of \( P \) at all then let \( o_1 \) be any point on the boundary of \( D_1 \). Note that there is some value of \( \epsilon > 0 \) such that \( D_1 \) is free to move by a distance of \( \epsilon \) in either of the two directions perpendicular to the direction \( \overrightarrow{p_1 o_1} \) while keeping \( D_1 \) in the interior of \( P \). However, movement in at least one of these directions will increase the distance \( |p_1 p_2| \), which is a contradiction since this distance was chosen to be maximal over all possible solutions to 2-disk.

Let \( x_1 \) and \( x_2 \) be two elementary objects of \( M(P) \). We define the distance between \( x_1 \) and \( x_2 \), denoted \( d(x_1, x_2) \) as \( 2r \), where \( r \) is the radius of the largest pair of equal-radius non-intersecting disks \( d_1 \) and \( d_2 \), contained in \( P \) and with \( d_i \) centered on \( x_i \), for \( i = 1, 2 \). There are two points to note about this definition of distance: (1) if the distance between two elementary objects is \( 2r \), then we can place two non-intersecting disks of radius \( r \) in \( P \), and (2) the distance from an elementary object to itself is not necessarily 0. Given two elementary objects it
is possible, in constant time, to compute the distance between them as well as the locations of 2 disks that produce this distance \[5\].

Let \( E \) be the set of elementary objects obtained by taking the union of the following three sets of elementary objects:

1. the set of vertices of \( M(P) \),
2. the set of elementary line segments obtained by splitting each straight line segment of \( M(P) \) into at most two elementary objects.
3. the set of elementary parabolic arcs obtained by splitting each parabolic arc of \( M(P) \) into at most two elementary objects.

We call the diameter of \( E \) the maximum distance between any pair \( x, y \in E \), where distance is defined as above. Now, it should be clear from Lemma \[1\] that 2-disk can be solved by finding a pair of elements in \( E \) whose distance is equal to the diameter of \( E \). \[1\]

Thus, all that remains is to devise an algorithm for finding the diameter of \( E \). Let \( m \) denote the cardinality of \( E \) and note that, initially, \( m = O(n) \). Motivated by Clarkson and Shor \[8\], we compute the diameter using the following algorithm. We begin by selecting a random element \( x \) from \( E \) and finding the element \( x' \in E \) whose distance from \( x \) is maximal, along with the corresponding radius \( r \). This can be done in \( O(m) \) time, since each distance computation between two elementary objects can be done in constant time. Note that \( r \) is a lower bound on \( r^* \). We use this lower bound to do trimming and pruning on the elements of \( E \).

We trim each element \( y \in E \) by partitioning \( y \) into two subarcs\[2\] each of which may be empty. The subarc \( y_\geq \) is the part of \( y \) supporting disks of radius greater than or equal to \( r \). The subarc \( y_< \) is the remainder of \( y \). We then trim \( y_< \) from \( y \) by removing \( y \) from \( E \) and replacing it with \( y_\geq \). During the trimming step we also remove from \( E \) any element that does not support a disk of radius greater than \( r \). Each such trimming operation can be done in constant time, resulting in an \( O(m) \) running time for this step.

Next, we prune \( E \). For any arc \( y \in E \), the lowest point of \( y \) is its closest point to the boundary of \( P \). In the case of ties, we take a point which is closest to one of the endpoints of \( P \). By the definition of elementary objects, the lowest point of \( y \) is therefore an endpoint of \( y \). The closed disk with radius \( r \) centered on the lowest point of \( y \) is denoted by \( D(y) \). We discard all the elements \( y \in E \) such that \( D(y) \cap D(x) \neq \emptyset \) for all \( x \in E \).

Pruning can be performed in \( O(m \log m) \) time by computing, for each lowest endpoint \( p \), a matching lowest endpoint \( q \) whose distance from \( p \) is maximal and then discarding \( p \) if \( |pq| \leq 2r \). This computation is known as all-pairs furthest neighbors and can be completed in \( O(m \log m) \) time \[1\].

Once all trimming and pruning is done, we have a new set of elementary objects \( E' \) on which we recurse. The recursion completes when \( |E'| \leq 2 \), at

\[1\] Here we use the term “pair” loosely, since the diameter may be defined by the distance between an elementary object and itself.
\[2\] We use the term subarc to mean both parts of segments and parts of parabolic arcs.
which point we compute the diameter of $E'$ in constant time using a brute-force algorithm. We output the largest pair of equal-radius non-overlapping disks found during any iteration of the algorithm.

To prove that this algorithm is correct we consider a pair of non-intersecting disks $D_1$ and $D_2$, each contained in $P$ and having radius $r^*$, centered at $p_1$ and $p_2$, respectively, such that the Euclidean distance $|p_1p_2|$ is maximal. The following lemma shows that $p_1$ and $p_2$ are not discarded from consideration until an equally good solution is found.

**Lemma 2.** If, during the execution of one round, $\{p_1, p_2\} \subset \cup E$ and $r < r^*$, then $\{p_1, p_2\} \subset \cup E'$ at the end of the round.

**Proof.** We need to show that at the end of the round, there exists elementary objects $y_1, y_2 \in E'$ such that $p_1 \in y_1$ and $p_2 \in y_2$. More specifically, we need to show there exists $y_1, y_2 \in E$ such that $p_1$, respectively $p_2$ is not trimmed from $y_1$, respectively $y_2$, and $y_1$ and $y_2$ are not pruned.

To see that $p_1$ and $p_2$ are not trimmed from any elementary object that contains them we simply note that $p_1$ and $p_2$ both support disks of radius $r^* > r$ and are therefore not trimmed.

To prove that $y_1$ and $y_2$ are not pruned we subdivide the plane into two open halfspaces $H_1$ and $H_2$ such that all points in $H_1$ are closer to $p_1$ than to $p_2$ and vice-versa. We denote by $L$ the line separating these two halfspaces.

Recall that, after trimming, an elementary object $x$ is only pruned if $D(x) \cap D(y) \neq \emptyset$ for all $y \in E$. We will show that $D(y_1) \subseteq H_1$ and $D(y_2) \subseteq H_2$, therefore $D(y_1) \cap D(y_2) = \emptyset$ and neither $y_1$ nor $y_2$ are pruned. It suffices to prove that $D(y_1) \subseteq H_1$ since a symmetric argument shows that $D(y_2) \subseteq H_2$. We consider three separate cases depending on the location of $p_1$ on $M(P)$.

**Case 1:** $p_1$ is a vertex of $M(P)$. In this case we choose $y_1$ to be the singleton elementary object $\{p_1\}$. Thus, $D(y_1)$ is centered at $p_1$. Furthermore, the distance between $p_1$ and $L$ is at least $r^* > r$. Therefore, one point of $D(y_1)$ is contained in $H_1$ and $D(y_1)$ does not intersect the boundary of $H_1$ so it must be that $D(y_1) \subseteq H_1$.

**Case 2:** $p_1$ lies in the interior of a straight line segment of $M(P)$. Let $p'_1$ be the lower endpoint of $y_1$. Let $\theta$ be the angle $(p_2p_1, p_1p'_1)$ (see Fig. 3). If $\theta \in [-\pi/2, \pi/2]$ then we can move $p_1$ slightly in the direction opposite to $p'_1$ while keeping $D_1$ inside $P$, thus contradicting the assumption that $|p_1p_2|$ is maximal. Therefore $\theta \in [\pi/2, 3\pi/2]$, which implies that $D(y_1)$ lies in $H_1$.

**Case 3:** $p_1$ lies in the interior of a parabolic arc of $M(P)$. In this case $D_1$ is tangent to an edge $e_1$ of $P$ and touches one of its vertices $v$. $p'_1$ still denotes the lower endpoint of $y_1$. Without loss of generality, assume that $e_1$ is parallel to the $x$-axis and $x(p'_1) < x(p_1)$ (see Fig. 4). Let $L'$ be the line parallel to $L$ that crosses the segment $[p_1, p_2]$ and that is tangent to $D_1$. We denote by $o_1$ the point where $e_1$ is tangent to $D_1$, and we denote by $o'_1$ the point such that $(o_1, o'_1)$ is a diameter of $D_1$. Finally, the convex hull of $D_1$ and $D(y_1)$ is denoted by $C$.

It must be that $x(p_2) > x(p_1)$, otherwise $p_1$ and $D_1$ could be moved in the positive $x$ direction while keeping $D_1$ in $P$. This would increase the distance $|p_1p_2|$ which is defined as maximal. It follows that $L'$ is tangent to $D_1$ along the
counterclockwise arc \((o'_1, o_1)\). Then \(L'\) is tangent to \(C\), so by convexity \(C\) lies on the same side of \(L'\) as \(p_1\) which implies that \(D_1\) is contained in \(H_1\).

Let \(d_i\) denote the distance of the furthest element in \(E\) from \(x_i\), and suppose for the sake of analysis that the elements of \(E\) are labeled \(x_1, \ldots, x_n\) so that \(d_i \leq d_{i+1}\). The following lemma helps to establish the running time of the algorithm.

**Lemma 3.** If we select \(x = x_i\) as the random element, then we discard all \(x_j \in E\) such that \(j \leq i\) from \(E\).
Proof. For any $j \leq i$, either $x_j$ does not support a disk of radius greater than $d_i$, or every point on $x_j$ that supports a disk of radius $d_i$ is of distance at most $d_i$ from any other point of $M(P)$ that supports a disk of radius $d_i$.

In the first case, $x_j$ is removed from $E$ by trimming. In the second case, $D(x_j) \cap D(x_k) \neq \emptyset$ for all $x_k \in E$ and $x_j$ is removed by pruning.

Finally, we state and prove our main theorem.

**Theorem 1.** The above algorithm solves 2-disk in $O(n \log n)$ expected time.

**Proof.** The algorithm is correct because, by Lemma 2, it never discards $p_1$ nor $p_2$ until it has found a solution with $r = r^*$, at which point it has already found an optimal solution that will be reported when the algorithm terminates.

To prove the running time of the algorithm, we use the following facts. Each round of the algorithm can be completed in $O(m \log m)$ time where $m$ is the cardinality of $E$ at the beginning of the round. By Lemma 3, when we select $x_i$ as our random element, all elements $x_j$ with $j \leq i$ disappear from $E$. Therefore, the expected running time of the algorithm is given by the recurrence

$$T(m) \leq \frac{1}{m} \sum_{i=1}^{m} T(m-i) + O(m \log m),$$

which readily solves to $O(m \log m)$. Since $m \in O(n)$, this completes the proof.

### 4 Conclusions

We have given a randomized algorithm for 2-disk that runs in $O(n \log n)$ expected time. The algorithm is considerably simpler than the $O(n \log^3 n)$ algorithm of Bespamyatnikh [4] and has the additional advantage of solving the more general problem of polygons with holes. Although we have described our algorithm as performing computations with distances, these can be replaced with squared distances to yield an algorithm that uses only algebraic computations.

In the algebraic decision tree model of computation, one can also prove an $\Omega(n \log n)$ lower bound on any algorithm for 2-disk through a reduction from max-gap [12]. Suppose that the input to max-gap is $y_1, \ldots, y_n$. Without loss of generality one can assume that $y_1 = \min\{y_i : 1 \leq i \leq n\}$ and $y_n = \max\{y_i : 1 \leq i \leq n\}$. We then construct a rectangle with top and bottom sides at $y_1$ and $y_n$, respectively, and with width $2(y_n - y_1)$. The interior of this rectangle is then partitioned into rectangles with horizontal line segments having $y$ coordinates $y_1, \ldots, y_n$. See Fig. 5 for an example.

It should then be clear that the solution to 2-disk for this problem corresponds to placing two disks in the rectangle corresponding to the gap between $y_i$ and $y_{i+1}$ which is maximal, i.e., it gives a solution to the original max-gap problem. Since this reduction can be easily accomplished in linear time and max-gap has an $\Omega(n \log n)$ lower bound, this yields an $\Omega(n \log n)$ lower bound on 2-disk.

The above reduction only works because we allow polygons with holes. An interesting open problem is that of determining the complexity of 2-disk when restricted to simple polygons. Is there a linear time algorithm?
Fig. 5. Reducing MAX-GAP to 2-DISK.

References

Maximum Red/Blue Interval Matching with Applications

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Abstract. In this paper, we consider the problem of computing a maximum cardinality matching among a set \( I \) of \( n \) intervals that are colored as either red or blue, such that a pair of intervals in \( I \) can be matched only if they overlap with each other and have different colors. This problem arises in some applications such as radiosurgery treatment planning. We present a greedy algorithm for this problem that runs in \( O(n \log \log n) \) time for sorted input. We also solve a useful generalization of this red/blue interval matching problem in the same time bound.

1 Introduction

Consider a set of \( n \) input intervals, \( I = \{I_1, I_2, \ldots, I_n\} \), on the \( x \)-axis that are colored as either red or blue, where each interval \( I_i = [le(i), re(i)] \) is specified by its left endpoint \( le(i) \) and right endpoint \( re(i) \), with \( le(i) < re(i) \). Two intervals \( I_i = [le(i), re(i)] \) and \( I_j = [le(j), re(j)] \) are disjoint to each other if \( re(i) < le(j) \) or \( re(j) < le(i) \); otherwise, they overlap. A graph \( G \) is called an interval graph if there exists a set \( I_G \) of intervals such that there is a one-to-one correspondence between the vertices of \( G \) and the intervals in \( I_G \) and such that any two vertices in \( G \) are connected by an edge if and only if their corresponding intervals in \( I_G \) overlap. Such an interval set \( I_G \) is called an interval model of \( G \). An interval graph \( G \) is said to be proper if and only if there is an interval model \( I_G \) of \( G \) such that no interval in \( I_G \) is contained within any other interval in \( I_G \). Interval graphs find applications in many areas, such as VLSI design, scheduling, biology, medicine, traffic control, and archeology \[5\]. In this paper, we assume that an interval model of the corresponding interval graph is already given (as this is often the case in applications). We will refer to interval \( I_i \), interval \( [le(i), re(i)] \), and vertex \( i \) (corresponding to interval \( i \)) interchangeably.

A matching in a graph \( G \) is a subset \( M \) of the edges of \( G \) such that no two distinct edges in \( M \) are incident to the same vertex. The problem of computing...
maximum matchings in graphs has many applications and has received a lot of attention [4].

In this paper, we consider the following matching problem on a set $I$ of $n$ red/blue intervals: Find a maximum cardinality matching $M$ in $I$ such that two intervals can be matched in $M$ only if they overlap and their colors are different. This problem, in fact, is that of computing a maximum cardinality matching in the interval graph $G$ of $I$ that is a bipartite graph [4].

The best known maximum bipartite matching algorithm on a general $n$-vertex graph is due to Hopcroft and Karp [5], which takes $O(n^{2.5})$ time. To our best knowledge, no algorithm for the specific matching problem on bipartite interval graphs was previously known. There are only a few known results for matching problems on interval graphs and some related graphs. In particular, Moitra and Johnson [7] gave an $O(n \log n)$ time algorithm for maximum matching in ordinary (non-bipartite) interval graphs. For maximum matching in the complement graph of an ordinary interval graph, Andrews and Lee [2] designed an $O(n \log n)$ time algorithm, and Andrews et al. [1] provided an $O(n)$ time algorithm for sorted input.

The maximum red/blue interval matching problem that we consider arises in the study of radiosurgery treatment planning. Radiosurgery is a non-invasive technique for eradicating localized benign and malignant tumors by using radiation beams. The shape of a radiation beam is formed by a device called collimator [12]. The distribution of radiation for a treatment is often prescribed by a so-called intensity map of radiation. An intensity map for a multileaf collimator based radiosurgery system [12] may consist of multiple tracks (called arcs). A pair of collimator leaves is assigned to each track. On every track, the prescribed radiation dose is often delivered as a number of (possibly overlapping) sub-tracks, each of which can be viewed as an interval on that track. For example, in Figure 1 we can treat $I_j$ (resp., $I_k$) as an interval on the track $T'$ (resp., $T''$). A radiosurgery system usually delivers a unit of radiation to a portion of a track as specified by one of its sub-tracks, with its two collimator leaves defining the two endpoints of such a sub-track. For two consecutive tracks, such as $T'$ and $T''$ in Figure 1 suppose that there are two sub-tracks, one on each track, whose intervals overlap (here, although the intervals are really on different tracks, we consider their overlapping as if the intervals were all on the $x$-axis). Then the two corresponding sub-tracks on the two consecutive tracks can be delivered by the radiosurgery system in the same time, thus shortening the total treatment time. (The overlapping among the sub-tracks on the same track does not help since they must be delivered one at a time.) Suppose we color the intervals on $T'$ (resp., $T''$) as red (resp., blue). Then one of the useful problems in treatment planning for intensity-modulated radiation therapy with multileaf collimation [3,12] is to obtain the maximum number of red/blue interval pairs such that the two intervals in each pair overlap. This gives rise to our maximum red/blue interval matching problem.

We also need to study a slightly more general version of the maximum red/blue interval matching problem. The endpoints of the sub-tracks (intervals)
in the above radiosurgery treatment planning setting are specified by pairs of leaves of a collimator. Due to some mechanical constraints (e.g., for preventing two opposite collimator leaves on two consecutive tracks from colliding into each other and being damaged), a system-dependent threshold value \( d > 0 \) is actually specified for the overlapping of any two red/blue intervals \([11]\). Precisely, the length of the overlapping portion of any two such intervals must be no smaller than the threshold value \( d \). Hence, the \textit{threshold overlapping maximum red/blue interval matching problem} is to find a maximum matching of \( n \) red/blue intervals subject to that two red/blue intervals can be matched only if the length of their overlapping portion is \( \geq d \). Obviously, this problem is a generalized version of the maximum red/blue interval matching problem.

We present an \( O(n \log \log n) \) time algorithm for the maximum red/blue interval matching problem if the endpoints of the input intervals are already given sorted; otherwise, the algorithm takes \( O(n \log n) \) time. Our algorithm is of a greedy nature, and is based on the plane sweeping technique \([8]\) and some integer data structures \([9,10]\). For the special case of a proper interval graph with sorted interval endpoints, we give an \( O(n) \) time matching algorithm. We also show an easy \( O(n) \) time reduction from the threshold overlapping maximum red/blue interval matching problem to the maximum red/blue interval matching problem for sorted input.

The rest of this paper is organized as follows. Section 2 gives some needed preliminaries. Section 3 presents our maximum red/blue interval matching algorithms. Section 4 shows the reduction from the threshold overlapping maximum red/blue interval matching problem to the maximum red/blue interval matching problem.

\section{Preliminaries}

The input consists of a set of \( n \) red or blue intervals, \( I = \{I_1, I_2, \ldots, I_n\} \). The endpoints of the intervals in \( I \) are arbitrary numbers on the \( x \)-axis. Each \( I_i \) is associated with a color \( C(I_i) \in \{\text{red, blue}\} \). To avoid cluttering the exposition, we assume without loss of generality (WLOG) that no two input intervals share the same endpoint (i.e., the \( 2n \) endpoints are distinct). Our algorithms can be easily modified for the general case.

We first sort the \( 2n \) endpoints of \( I \) from left to right if they are not given sorted. The sorting takes \( O(n \log n) \) time. From now on, unless otherwise spec-
ified, we assume that the $2n$ endpoints of $I$ are available in this sorted order and are stored in an array $A$. Further, we treat each interval endpoint of $I$ as its rank in this sorted array $A$, i.e., we simply treat each interval endpoint as an integer in $\{1, 2, \ldots, 2n\}$. We also assume WLOG that the intervals in $I$ have been relabeled such that $i < j$ implies that $re(i)$ occurs before $re(j)$ in the sorted array $A$ of endpoints.

If the input interval graph is proper, then we assume that no interval in $I$ contains another interval in $I$.

Given a matching $M$ in $I$, we say that an interval $i$ is in $M$ if $i$ is matched by $M$. An interval is free if it is unmatched with respect to $M$. Interval $i$ matching with $j$ in $M$ is denoted as $mate_M(i) = j$. A matched pair of intervals $i$ and $j$ in $M$ is denoted as $match_M(i, j)$.

It is known that when the keys of a priority queue are all from a universe $\{1, 2, \ldots, m\}$, there is a data structure which supports each of the following operations in $O(\log \log m)$ time \cite{9,10}: Insert, Delete, Member, Empty, Min, Max, Predecessor, and Successor. Our algorithm makes use of such integer priority queues (with $m = 2n$).

## 3 The Maximum Red/Blue Matching Algorithms

Our maximum red/blue interval matching algorithm is based on the following key observation.

**Lemma 1.** Let $J$ be the interval in $I$ such that its right endpoint $re(J)$ is the smallest among the right endpoints of all intervals in $I$. WLOG, let $C(J) = \text{red}$ (the case of $C(J) = \text{blue}$ is symmetric). Then the following holds.

(i) If $J$ does not overlap with any blue interval in $I$, then $J$ is free in any matching $M$ in $I$.

(ii) If $J$ overlaps with any blue interval $J'$ in $I$, then $re(J) < re(J')$. Further, let $J''$ be the blue interval in $I$ that overlaps with $J$ and whose right endpoint $re(J'')$ is the smallest among all blue intervals of $I$ that overlap with $J$. Then for any maximum matching $M^*$ in $I$, there is another maximum matching $M'$ in $I$ such that $mate_{M'}(J) = J''$.

**Proof:** Since the other statements of this lemma are easy to show, we only prove that for any maximum matching $M^*$ in $I$, there is another maximum matching $M'$ in $I$ such that $mate_{M'}(J) = J''$. Let $M^*$ be any maximum matching in $I$. Note that it is impossible that both $J$ and $J''$ are free with respect to $M^*$ (otherwise, we can add $(J, J'')$ to $M^*$, a contradiction to the optimality of $M^*$). WLOG, we assume that the pair $(J, J'') \notin M^*$ (otherwise, $M' = M^*$). We consider three cases: (1) $J$ is free and $J''$ is in $M^*$, (2) $J$ is in $M^*$ and $J''$ is free, and (3) both $J$ and $J''$ are in $M^*$ (but $(J, J'') \notin M^*$). We will show that in each case, there is another maximum matching $M'$ in $I$ such that $mate_{M'}(J) = J''$.

It is easy to obtain a maximum matching $M'$ in $I$ from $M^*$ for Cases (1) and (2), as follows: Replace $mate_{M^*}(J', J'')$ by $mate_{M^*}(J, J'')$ in Case (1), and replace $mate_{M^*}(J, J')$ by $mate_{M^*}(J, J'')$ in Case (2), where $J' \notin \{J, J''\}$.
is the interval (with an appropriate color) that is matched with \( J'' \) or \( J \) in \( M^* \) in these two cases, respectively; let \( M' \) keep the other matched pairs of \( M^* \) unchanged.

![Fig. 2.](image)

For Case (3), let \( \text{mate}_{M^*}(J) = J_b \) and \( \text{mate}_{M^*}(J'') = J_r \), where \( C(J_b) = \text{blue} \) and \( C(J_r) = \text{red} \). We aim to prove that \( J_b \) and \( J_r \) overlap (and hence can be matched). By the definitions of \( J \) and \( J'' \), we have \( \text{re}(J) < \text{re}(J'') < \text{re}(J_b) \) (see Figure 2). This and the fact that \( J \) and \( J_b \) overlap imply that the interval \([\text{re}(J), \text{re}(J'')]\) on the \( x \)-axis is completely contained by \( J_b \). Also, by the definition of \( J_r \), we have \( \text{re}(J) < \text{re}(J_b) < \text{re}(J_r) < \text{re}(J'') \) (see Figure 2(a)) or \( \text{re}(J) < \text{re}(J'') < \text{re}(J_r) \) (see Figure 2(b)) holds. In either situation, due to the fact that \( J'' \) and \( J_r \) overlap, \( J_r \) must overlap with the interval \([\text{re}(J), \text{re}(J'')]\). Hence, \( J_r \) and \( J_b \) must overlap (precisely, on at least one endpoint of the interval \([\text{re}(J), \text{re}(J'')]\)), and can be matched. Therefore, we can obtain a maximum matching \( M' \) in \( I \) from \( M^* \) by replacing \( \text{match}_{M^*}(J, J_b) \) and \( \text{match}_{M^*}(J'', J_r) \) by \( \text{match}_{M'}(J, J_b) \) and \( \text{match}_{M'}(J_r, J'') \), and letting \( M' \) keep the other matched pairs of \( M^* \) unchanged. This finishes the proof of this lemma.

Lemma 1 immediately implies the following greedy algorithm: Find intervals \( J \) and \( J'' \) in \( I \) as defined in Lemma 1 let them form a matched pair, and solve the problem recursively on the remaining intervals in \( I - \{J, J''\} \). Of course, an efficient way to carry out this greedy method is by using a non-recursive algorithm. Indeed, our greedy algorithm is based on the plane sweeping technique [8] and integer priority queues [9,10].

We first give an overview of our algorithm. It uses a sweep line to scan the endpoints of the intervals of \( I \) in the increasing order. It maintains two integer priority queues, \( Q_R \) (for the red intervals) and \( Q_B \) (for the blue intervals), such that all intervals in \( Q_R \) or \( Q_B \) are currently being intersected by the sweep line. Note that if at the same time, a red (resp., blue) interval \( J' \) (resp., \( J'' \)) is stored in \( Q_R \) (resp., \( Q_B \)), then \( J' \) and \( J'' \) overlap and are possible to form a matched pair. An interval \( J \) is inserted into (resp., deleted from) the corresponding priority queue when its left (resp., right) endpoint \( \text{le}(J) \) (resp., \( \text{re}(J) \)) is encountered by the sweep line.
The details of our algorithm follow next.

**Algorithm Max-RB-Matching** ($I$);

1. Initialize the integer priority queues $Q_R$ and $Q_B$ as containing no key values. Let $L$ be the sorted list of all endpoints of $I$ in the increasing order. Let the output maximum matching in $I$ initially be $M^* = \phi$. Mark every interval of $I$ as “free”.

2. While $L$ is not empty do
   (a) Remove an endpoint $p$ from the head of $L$. Let $J$ be the interval of $I$ such that $p$ is an endpoint of $J$. WLOG, assume that $C(J) = \text{red}$ (the other case is symmetric).
   (b) If $p$ is the left endpoint $le(J)$ of $J$, then insert the right endpoint $re(J)$ of $J$ into the priority queue $Q_R$. (Note that $re(J)$ is not removed from the list $L$.) Go to the next iteration.
   (c) If $J$ is already in $M^*$, then go to the next iteration. Else, $p$ is the right endpoint $re(J)$ of $J$. Search in $Q_B$ for the immediate successor $\text{Succ}_{Q_B}(re(J))$ of $re(J)$ in $Q_B$.
      If $\text{Succ}_{Q_B}(re(J))$ does not exist in $Q_B$, then $Q_B$ contains no blue interval. Delete $re(J)$ from $Q_R$. (The interval $J$ is not matched.)
      If $\text{Succ}_{Q_B}(re(J))$ does exist in $Q_B$, then let it be the right endpoint $re(J'')$ of a blue interval $J''$. Add a matched pair $(J, J'')$ to the output maximum matching $M^*$. Delete $re(J)$ from $Q_R$ and $re(J'')$ from $Q_B$. Mark both $J$ and $J''$ as “matched”.

End of the while loop.

**Lemma 2.** Algorithm Max-RB-Matching computes a maximum matching for $n$ red and blue intervals in $O(n \log \log n)$ time, provided that the endpoints of the intervals are already given sorted.

**Proof:** The correctness of the algorithm follows from Lemma 1. WLOG, we only consider the situation when the right endpoint $re(J)$ of a red interval $J$ is encountered by the sweep line, and further, we assume that $J$ is currently not in $M^*$. Note that all blue intervals of $I$ that are currently being intersected by the sweep line and that are not in $M^*$ are stored in the priority queue $Q_B$. If no blue interval is in $Q_B$ (as found by the search of $\text{Succ}_{Q_B}(re(J))$ in $Q_B$), then by Lemma 1, $J$ should be free. If the immediate successor $\text{Succ}_{Q_B}(re(J))$ of $re(J)$ in $Q_B$ is obtained, then again by Lemma 1 $(J, J'')$ should be matched in $M^*$, where $J''$ is the blue interval having $\text{Succ}_{Q_B}(re(J))$ as its right endpoint.

Analyzing the time complexity of Algorithm Max-RB-Matching is easy. Each of the $n$ right endpoints of the $n$ input intervals is inserted once into an integer priority queue and deleted once from such a queue, and it gives rise to at most one Successor search in such a queue. Since each of the Insert, Delete, and Successor operations takes $O(\log \log n)$ time \cite{9,10}, the total running time of the algorithm is $O(n \log \log n)$.

We now consider the proper interval graph case (no interval in $I$ contains another interval in $I$).
Lemma 3. A maximum matching for a set $I$ of $n$ red and blue intervals can be computed in $O(n)$ time, provided that the endpoints of the intervals in $I$ are already given sorted and no interval in $I$ contains another interval in $I$.

Proof: This follows quite easily from Lemma 1. We maintain two sorted lists $I_R$ and $I_B$, with $I_R$ (resp., $I_B$) containing the red (resp., blue) intervals in $I$ ($I_R \cup I_B = I$). Note that since no interval in $I$ contains another interval in $I$, we can order the intervals in $I_R$ and $I_B$ based only on their left endpoints.

Our $O(n)$ time algorithm for computing a maximum matching in $I$ is very similar to merging two sorted lists. We consider the first element $J$ of $I_R$ and the first element $J'$ of $I_B$. If $J$ and $J'$ overlap, then by Lemma 1, we match $J$ and $J'$ and remove $J$ from $I_R$ and remove $J'$ from $I_B$. Otherwise, $J$ and $J'$ do not overlap with each other. Assume WLOG that $le(J) < le(J')$. Then by Lemma 1 again, we remove $J$ from $I_R$ (but $J'$ stays in $I_B$). In either case, we repeat the matching process on the remaining elements of $I_R$ and $I_B$, until one of $I_R$ or $I_B$ becomes empty. It is easy to see that since $I_R$ and $I_B$ are both sorted, this algorithm takes $O(n)$ time. 

\[\square\]

4 Threshold Overlapping Maximum Red/Blue Matching

As a generalization of the maximum red/blue interval matching problem, we solve the threshold overlapping maximum red/blue interval matching problem by reducing it to the maximum red/blue interval matching problem. Again, we assume that the endpoints of $I$ are already given sorted. However, at this section, we need to use the original input values of the endpoints, not their integral ranks as in Section 3.

Let $d > 0$ be the specified threshold value for the minimum acceptable length of overlapping between any two red/blue intervals. The following simple observation is a key to our reduction.

Lemma 4. For any red (resp., blue) interval $J_r$ (resp., $J_b$), let $J_r(d)$ (resp., $J_b(d)$) be the interval obtained from $J_r$ (resp., $J_b$) by cutting a length of $d/2$ away from each end of $J_r$ (resp., $J_b$). Then $J_r$ and $J_b$ have an overlapping length $\geq d$ if and only if $J_r(d)$ and $J_b(d)$ overlap (see Figure 3).

![Fig. 3](image-url)
Proof: Easy and omitted.

Note that if an interval $J$ of $I$ has a length $< d$, then applying to $J$ the cut-away operation as defined in Lemma 4 creates an empty interval $J(d)$ that we should throw away from further consideration (since $J$ will never overlap with any interval in a length $\geq d$). We say that such an interval $J$ is short. In fact, all short intervals of $I$ are not good and should be identified, because the distances between their endpoints are less than the threshold value $d$. (Thus, they possibly cause the opposite collimator leaves on the same tracks, as well as on consecutive tracks, that define the endpoints of these and other intervals to collide into each other.) Hence, all short intervals of $I$ are not useful medically and should be removed even before one considers solving the threshold overlapping maximum red/blue interval matching problem. WLOG, we assume that there are no short intervals in $I$. Let $I(d)$ be the set of intervals obtained from $I$ by applying the cut-away operation of Lemma 4.

All that is left for us to do is to put the endpoints of $I(d)$ into a sorted order. Note that the cut-away operations may cause the order of the endpoints of $I$ to be different from the order of the endpoints of $I(d)$, and we want to sort the endpoints of $I(d)$ in $O(n)$ time.

Let $LE(I)$ (resp., $LE(I(d))$) be the list of left endpoints of $I$ (resp., $I(d)$), and $RE(I)$ (resp., $RE(I(d))$) be the list of right endpoints of $I$ (resp., $I(d)$). Note that both $LE(I)$ and $RE(I)$ are already in a sorted order. Since the cut-away operations basically shift all points in $LE(I)$ (resp., $RE(I)$) to the right (resp., left) by the same distance of $d/2$, $LE(I(d))$ and $RE(I(d))$ are in the same sorted order as $LE(I)$ and $RE(I)$, respectively. Hence, to create a sorted list for all endpoints of $I(d)$, we simply merge $LE(I(d))$ and $RE(I(d))$, in $O(n)$ time. This finishes our reduction from the threshold overlapping maximum red/blue interval matching problem to the maximum red/blue interval matching problem.

**Lemma 5.** In $O(n)$ time, the threshold overlapping maximum red/blue interval matching problem can be reduced to the maximum red/blue interval matching problem, if the endpoints of the $n$ input intervals are already given in a sorted order.

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References

Computing Farthest Neighbors
on a Convex Polytope*

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Abstract. Let \( N \) be a set of \( n \) points in convex position in \( \mathbb{R}^3 \). The
farthest-point Voronoi diagram of \( N \) partitions \( \mathbb{R}^3 \) into \( n \) convex cells.
We consider the intersection \( G(N) \) of the diagram with the boundary of
the convex hull of \( N \). We give an algorithm that computes an implicit
representation of \( G(N) \) in expected \( O(n \log^2 n) \) time. More precisely,
we compute the combinatorial structure of \( G(N) \), the coordinates of its
vertices, and the equation of the plane defining each edge of \( G(N) \). The
algorithm allows us to solve the all-pairs farthest neighbor problem for \( N \)
in expected time \( O(n \log^2 n) \), and to perform farthest-neighbor queries
on \( N \) in \( O(\log^2 n) \) time with high probability. This can be applied to find
a Euclidean maximum spanning tree and a diameter 2-clustering of \( N \)
in expected \( O(n \log^4 n) \) time.

1 Introduction

Let \( N \) be a set of \( n \) points in three dimensions. Its diameter is the maximum
distance between any two points in \( N \). The problem of computing the diame-
ter has been intensively studied in the past two decades. Indeed, back in 1985
After Clarkson and Shor [4] gave a simple randomized algorithm that runs in
optimal expected \( O(n \log n) \) time, most work on the problem has concentrated on
finding a matching deterministic algorithm. After considerable efforts by several
that run in \( O(n \log^2 n) \) time. Finally, Ramos [8] solved the problem in optimal
time \( O(n \log n) \).

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The all-pairs farthest neighbors problem for a set \( N \) of \( n \) points in three dimensions is to compute, for each point \( p \) in \( N \), the point of \( N \) farthest from \( p \). This natural generalization of the diameter problem has several applications [1]. While all-pairs nearest neighbors in fixed dimension \( d \) can be computed in optimal \( O(n \log n) \) time [9], no algorithm with similar efficiency is known for the all-pairs farthest neighbors. Agarwal et al. [1] showed that 3-dimensional all-pairs farthest neighbors can be computed in \( O(n^{4/3} \log^{4/3} n) \) time, and pose closing the gap between this and the only lower bound of \( \Omega(n \log n) \) as a challenging open problem.

Our work is motivated by the fact that two-dimensional all-pairs farthest neighbors can be computed in linear time if the points are the vertices of a given convex polygon, even though the problem has complexity \( \Omega(n \log n) \) for arbitrary points [2]. We consider the all-pairs farthest neighbor problem for a set of \( n \) points that form the vertices of a convex polytope in 3 dimensions, and show that it can be solved in expected \( O(n \log n) \) time. This is the first algorithm on a restricted all-pairs farthest neighbor problem in three dimensions that runs in near-linear time.

Our algorithm computes a data structure that is interesting in its own right: given a convex polytope \( P \) with \( m \) vertices, and \( n \) sites lying on its surface, we compute an implicit representation of the intersection of the farthest-point Voronoi diagram of these sites with the surface of \( P \) in expected \( O(m \log m + n \log n \log m) \) time.

Our result can be used to solve the following problems, as a direct application of Agarwal et al.’s results [1]:

- Bi-chromatic farthest neighbors: Given a set \( R \) of \( n \) “red” points and another set \( B \) of \( m \) “blue” points in 3 dimensions such that \( R \cup B \) is in convex position, we find for each red point \( r \in R \) the farthest blue point from \( r \) in expected \( O((n + m) \log^2 (n + m)) \) time.
- External farthest neighbors: Given a set \( N \) of \( n \) points in 3 dimension in convex position and its partition \( N_1, N_2, \ldots, N_m \) into \( m \) subsets, we compute in expected \( O(n \log^3 n) \) time for each point \( p \) in \( N \), a farthest point in \( N \setminus N_i \), where \( p \in N_i \).
- Euclidean maximum spanning tree: Given a set \( N \) of \( n \) points in 3 dimensions in convex position, we compute in expected \( O(n \log^4 n) \) time a spanning tree of \( N \) whose edges have the maximum total length among all spanning trees, where the length of an edge is the Euclidean distance between its endpoints. From this tree we can compute a minimum diameter 2-clustering of \( N \) in linear time.

In the remainder of the paper, we consider only the original problem of computing all-pairs farthest neighbors for a set of points in convex position.

2 Preliminaries

Given a set \( N \) of points sites in \( \mathbb{R}^3 \) and a point site \( s \) not necessarily in \( N \), we define the (farthest-point) Voronoi cell \( \text{Vor}(s\mid N) \) of \( s \) with respect to \( N \) as the
set of points \( x \in \mathbb{R}^3 \) such that the Euclidean distance \( d(x,s) \) is larger than the distance \( d(x,s') \) to any site \( s' \in N \) with \( s' \neq s \). Voronoi cells are convex, and may be empty. The (farthest-point) Voronoi diagram of \( N \) is the partition of \( \mathbb{R}^3 \) into the Voronoi cells \( \text{Vor}(s|N) \), for \( s \in N \).

Let now \( P \) be the boundary of a convex polytope in three dimensions. Let \( N \) be a set of point sites lying on \( P \), and \( s \) a site on \( P \) not necessarily in \( N \). The Voronoi cell \( \text{Vor}(s|N) \) intersects \( P \) in a two-dimensional, possibly empty, Voronoi face \( \text{Vor}_P(s|N) \). Our first observation is that Voronoi faces are simply connected.

**Lemma 1.** Let \( P \) be the boundary of a 3-dimensional polytope, \( N \) a set of point sites on \( P \), and \( s \in N \). The Voronoi face \( \text{Vor}_P(s|N) \) is simply connected, that is, its boundary is a simple closed curve.

**Proof.** Let \( p, q \) be two points in \( \text{Vor}_P(s|N) \). Let \( C_s(pq) \) be the two-dimensional cone with apex \( s \) spanned by \( pq \), and let \( L_s(pq) \) be the intersection \( C_s(pq) \cap P \). \( L_s(pq) \) is a path on \( P \) connecting \( p \) and \( q \). We prove that \( L_s(pq) \) lies entirely in \( \text{Vor}_P(s|N) \).

In fact, let \( x \in pq \). Since \( x \in \text{Vor}(s|N) \), \( N \) lies in the sphere \( S \) with center \( x \) and passing through \( s \). If we enlarge \( S \) by moving its center along the ray \( sx \) and keeping \( s \) on the sphere, \( N \) will remain inside the enlarged sphere. It follows that the entire portion of \( C_s(pq) \) not in the triangle \( spq \) is contained in \( \text{Vor}(s|N) \). Since \( L_s(pq) \) lies in this portion, we have \( L_s(pq) \subset \text{Vor}_P(s|N) \).

Moreover, if we centrally project \( L_s(pq) \) from \( s \), the result is a line segment. Since for any two points \( p, q \) in \( \text{Vor}_P(s|N) \) we have \( L_s(pq) \subset \text{Vor}_P(s|N) \), this implies that the central projection of \( \text{Vor}_P(s|N) \) from \( s \) is convex, and therefore simply connected.

It follows that a set of sites \( N \) on \( P \) partitions \( P \) into simply connected faces, defining a planar graph that we denote as \( G(N) = G_P(N) \). A face of \( G(N) \) is a Voronoi face, a vertex of \( G(N) \) is a point of equal distance from three sites, and therefore the intersection of an edge of the three-dimensional Voronoi diagram with \( P \). There can be at most two vertices defined by the same three sites (and this case can indeed arise). An edge of \( G(N) \) separates two Voronoi faces \( \text{Vor}_P(s|N) \) and \( \text{Vor}_P(s'|N) \), and therefore lies on the bisecting plane of the sites \( s \) and \( s' \).

**Theorem 1.** Let \( N \) be a set of \( n \) sites on a polytope \( P \). Then \( G(N) = G_P(N) \) has \( O(n) \) vertices, edges, and faces.

**Proof.** It follows from Lemma 1 that the number of faces is at most \( n \). A vertex has degree at least three. Euler’s formula now implies the linear bound on the number of vertices and edges.

The embedding of an edge \( e \) of \( G(N) \) in \( P \) is a polyline whose vertices are the intersections between the embedding of \( e \) and the edges of \( P \). If \( P \) has \( m \) edges, the embedding of \( e \) consists of at most \( m \) segments. The overall complexity of
Voronoi region of \( s \) on the top disk
\[ \frac{n}{2} \text{ points} \]
\[ \frac{n}{2} \text{ points} \]

**Fig. 1.** An example of \( G(N) \) with complexity of \( \Omega(n^2) \). \( P \) is the convex hull of \( N \). (a) If one puts \( n/2 \) points equidistantly on the boundary of the bottom disk, then their Voronoi cells partition the top disk equally. (b) Add a convex roof slightly above the top disk which consists of the remaining \( n/2 \) points. The number of intersections of the Voronoi cells of the points on the bottom disk with edges of the roof becomes \( \Omega(n^2) \). The fat edge of the roof intersects the shaded Voronoi cells.

the embedding of \( G(N) \) is therefore \( O(nm) \). This bound is tight, as the example of the modified \( n \)-Camembert in Fig. 1 shows.

To achieve subquadratic time, we cannot work with the explicit embedding of \( G(N) \) into \( P \). Instead, we will use a linear size representation of \( G(N) \). The representation stores the adjacency relations between vertices, edges, and faces of \( G(N) \). By Theorem 1 this has linear complexity. In addition, we record for each face the defining site, and for each vertex the coordinates of its embedding into \( P \).

We need the following lemma.

**Lemma 2.** Let \( N \) be a set of sites on a polytope \( P \), and let \( v \) be a vertex of \( G(N) = G_P(N) \). If the sites defining the faces around \( v \) are known in the correct order, then the coordinates of \( v \)'s embedding into \( P \) can be computed using a single ray shooting query on \( P \).

**Proof.** The sites define a line of equal distance, which intersects \( P \) in two points. The two points differ in the order of sites. The order of sites thus orients the line, and we can find the correct intersection point with \( P \) using a ray shooting query from infinity.

As a final observation, note that this approach to computing the farthest-point Voronoi diagram would fail if we tried to apply it to the nearest-point Voronoi diagram instead. In fact, the intersection of the nearest-point Voronoi diagram of \( n \) sites \( N \) with the boundary of the convex hull of \( N \) can have a quadratic number of faces; such an example can be easily designed.

In the following we will assume \( N \) to be in general position, which means that no five sites lie on a sphere and no four points lie on a circle.
3 The Strategy to Compute $G(N)$

Given a convex polytope $\mathcal{P}$ with $m$ vertices and a set $N$ of $n$ point sites on the surface $\mathcal{P}$, we show how to compute $G(N) = G_{\mathcal{P}}(N)$ in expected time $O(m \log m + n \log n \log m)$.

The first step of the algorithm is to compute, in $O(m \log m)$ time, a data structure that permits ray shooting queries in $\mathcal{P}$ with query time $O(\log m)$.

We then compute $G(N)$ by randomized incremental construction. We choose a random permutation $s_1, s_2, s_3, \ldots, s_n$ of $N$, and insert the sites in this order.

Let $N^i = \{s_1, s_2, \ldots, s_i\}$ be the set of the first $i$ inserted sites. The algorithm maintains the implicit representation of $G(N^i)$ while sites are added one by one, resulting in $G(N^n) = G(N)$. Note that the polytope $\mathcal{P}$ defining $G(N^i)$ does not change during the course of the algorithm.

Our algorithm is mostly a straightforward implementation of the randomized incremental paradigm using a conflict graph, and most of the lemmas below are analogous to those proven, say, in Mulumley’s book [5]. We do, however, need to cope with an unpleasant aspect of our diagram concerning the maintenance of the conflict graph. As we will see, conflicts can “jump” to another edge. We need $O(\log m)$ time to check a conflict between a site and an edge, which results in overall $O(n \log n \log m)$ expected time for the computation of $G(N)$.

4 Conflicts

A vertex $v$ of $G(N^i)$ is said to be in conflict with a site $s \in N \setminus N^i$ if $s$ is farther from $v$ than any of the sites that define $v$, that is, the sites whose faces are adjacent to $v$. This is equivalent to $v \in \text{Vor}_\mathcal{P}(s|N^i)$. Similarly, an edge $e$ of $G(N^i)$ is said to be in conflict with a site $s \in N \setminus N^i$ if $e \cap \text{Vor}_\mathcal{P}(s|N^i) \neq \emptyset$.

In addition to a representation of $G(N^i)$, our algorithm maintains a conflict list: for each not-yet-inserted site $s \in N \setminus N^i$, we keep a bidirectional pointer to a single vertex $X(s)$ of $G(N^i)$ in conflict with $s$. If no vertex of $G(N^i)$ is in conflict with $s$, we set $X(s)$ to a single edge of $G(N^i)$ in conflict with $s$. If no edge of $G(N^i)$ is in conflict with $s$ either, then $X(s) := \emptyset$.

Lemma 3. Let $s \in N \setminus N^i$. If $\text{Vor}_\mathcal{P}(s|N^i)$ is not empty, then the vertices and edges of $G(N^i)$ in conflict with $s$ form a connected subgraph of $G(N^i)$.

Proof. Suppose the vertices and edges are not connected. Then we can separate them using a curve $\gamma$ contained in $\text{Vor}_\mathcal{P}(s|N^i)$ that cuts $\text{Vor}_\mathcal{P}(s|N^i)$ into two non-empty components without intersecting vertices or edges of $G(N^i)$. This means that $\gamma$ is entirely contained in a face $\text{Vor}_\mathcal{P}(s'|N^i)$ of $G(N^i)$. Since $\text{Vor}_\mathcal{P}(s'|N^i)$ is simply connected by Lemma [1], $\gamma$ cannot be a closed curve. The endpoints of $\gamma$ lie in $\text{Vor}_\mathcal{P}(s'|N^i) \setminus \text{Vor}_\mathcal{P}(s|N^i) = \text{Vor}_\mathcal{P}(s'|N^i \cup \{s\})$. Since this set is connected by Lemma [1], there is a path $\gamma'$ connecting the endpoints of $\gamma$ through $\text{Vor}_\mathcal{P}(s'|N^i) \setminus \text{Vor}_\mathcal{P}(s|N^i)$. The concatenation of $\gamma$ and $\gamma'$ is a closed curve contained in $\text{Vor}_\mathcal{P}(s'|N^i)$. It separates the edges and vertices of $G(N^i)$, a contradiction to Lemma [1].
We consider now the insertion of the $i+1$st site $s_{i+1}$ into the data structure storing $G(N^i)$.

**Lemma 4.** During the insertion of $s_{i+1}$, one of the three following situations occurs (see Fig. 3):

(i) If $s_{i+1}$ has no conflicting vertex and no conflicting edge in $G(N^i)$, then $\text{Vor}_P(s_{i+1}|N^i) = \emptyset$ and $G(N^{i+1}) = G(N^i)$.
(ii) If $s_{i+1}$ has only one conflicting edge $e$ and no conflicting vertex in $G(N^i)$, then $\text{Vor}_P(s_{i+1}|N^i) \cap G(N^i)$ is a connected portion $pq$ of $e$.
(iii) If $s_{i+1}$ has at least one conflicting vertex in $G(N^i)$, then at least one endpoint of each edge in conflict with $s_{i+1}$ is also in conflict with $s_{i+1}$.

**Proof.** If $s_{i+1}$ has no conflicting vertex and no conflicting edge in $G(N^i)$, then $\text{Vor}_P(s_{i+1}|N^i)$ must be empty; otherwise, $\text{Vor}_P(s_{i+1}|N^i)$ would lie entirely within some face of $G(N^i)$, which is impossible since the face is simply connected in $G(N^{i+1})$. Thus if $\text{Vor}_P(s_{i+1}|N^i) \neq \emptyset$, then there must be a vertex or an edge of $G(N^i)$ in conflict with $s_{i+1}$.

Suppose now that an edge of $G(N^i)$ conflicts with $s_{i+1}$, but that none of its endpoints conflicts with $s_{i+1}$. Then Lemma 3 directly implies that no other vertex or edge of $G(N^i)$ is in conflict with $s_{i+1}$.

The third case trivially holds because the vertices and edges in conflict with $s_{i+1}$ form a connected subgraph of $G(N^i)$ by Lemma 3.

![Fig. 2.](image)

(a) $G(N^3)$. No vertex of $G(N^3)$ is in conflict with $s_4$, but edge $e$ conflicts with $s_4$. (b) $G(N^4)$.

The algorithm implements the three cases of Lemma 4. Consider the insertion of $s_{i+1}$ into the data structure for $G(N^i)$. We first follow the conflict pointer of $s_{i+1}$ to find the conflict $X(s_{i+1}) \in G(N^i)$ in constant time.

If $X(s_{i+1}) = \emptyset$, we have case (i) of the lemma, and nothing needs to be done.

If $X(s_{i+1})$ is an edge, no vertex of $G(N^i)$ is in conflict with $s_{i+1}$. We have therefore case (ii) of the lemma, and there is no other conflict of $s_{i+1}$ at all. We
can update our data structure to represent $G(N^{i+1})$ by removing a portion of $e$ and replacing it with an eye-like subgraph that consists of two edges induced by the two bisector planes between $s_{i+1}$ and each of the two sites defining $e$ in $N^i$. See Fig. 3. Updating the adjacency relations takes constant time. However, we also need to compute the coordinates of the two new vertices. By Lemma 2, this can be done by two ray shooting queries on $P$ in time $O(\log m)$.

If $X(s_{i+1})$ is a vertex, we have case (iii) of the lemma. By Lemma 3, the portion of $G(N^i)$ lying inside $\text{Vor}_P(s_{i+1}|N^i)$ is a connected subgraph $G$ of $G(N^i)$. We identify $G$ by traversing $G(N^i)$, as we only need to test conflicts between vertices and $s_{i+1}$, which takes constant time per test. The extremal edges of $G$ lie on the boundary of the new face $\text{Vor}_P(s_{i+1}|N^i)$. We shorten these edges by creating new vertices. After generating the new boundary by connecting these vertices, we finally delete $G$. If the complexity of $G$ is $k$, all this can be done in time $O(k \log m)$ using Lemma 2. See Fig. 3.

![Fig. 3. Illustrating two cases of Lemma 4](image)

## 5 Conflict Maintenance

It remains to see how we update the conflict list during the insertion of a new site $s_{i+1}$. Recall that we maintain for every site $s \in N \setminus N^i$ a bidirectional pointer to a vertex or edge of $G(N^i)$ in conflict with $s$, if there is one. More precisely, if there are vertices of $G(N^i)$ in conflict with $s$, then $X(s)$ is one of them. If no vertex of $G(N^i)$ is in conflict with $s$ and $\text{Vor}_P(s|N^i) \neq \emptyset$, then there must be an edge $e \subset G(N^i)$ in conflict with $s$ by Lemma 4 (ii). In this case, we set $X(s) = e$. Notice that if $X(s)$ is an edge, then no vertex in $G(N^i)$ conflicts with $s$ (see Lemma 4 (ii)).

When inserting $s_{i+1}$, a new face $\text{Vor}_P(s_{i+1}|N^i)$ is defined. If $\text{Vor}_P(s_{i+1}|N^i) \neq \emptyset$, then the vertices and edges in conflict with $s_{i+1}$ will be destroyed in $G(N^{i+1})$. These are exactly the vertices and edges of $G(N^i)$ that intersect $\text{Vor}_P(s_{i+1}|N^i)$. 

![Fig. 3. Illustrating two cases of Lemma 4](image)
Let $s$ be a non-inserted site in $N \setminus N^{i+1}$. If $X(s)$ intersects $\text{Vor}_P(s_{i+1}|N^i)$, then $X(s)$ is not defined in $G(N^{i+1})$ any more, and we need to update $X(s)$ by finding a vertex or edge of $G(N^{i+1})$ in conflict with $s$. Otherwise, namely if $X(S)$ does not intersect $\text{Vor}_P(s_{i+1}|N^i)$, then we do not need to update $X(s)$. The next lemmas will be used for such conflict update. The boundary of a set $R$ is denoted by $\partial R$.

**Lemma 5.** Suppose that $s \in N \setminus N^{i+1}$ and $X(s) \cap \text{Vor}_P(s_{i+1}|N^i) \neq \emptyset$. If $\text{Vor}_P(s|N^i) \subset \text{Vor}_P(s_{i+1}|N^i)$ then $\text{Vor}_P(s|N^i) = \emptyset$. Otherwise $\text{Vor}_P(s|N^i) \cap \partial \text{Vor}_P(s_{i+1}|N^i) \subset \text{Vor}_P(s|N^i+1)$.

**Proof.** If $\text{Vor}_P(s|N^i) \subset \text{Vor}_P(s_{i+1}|N^i)$, then all the vertices and edges of $G(N^i)$ lying inside $\text{Vor}_P(s|N^i)$ would be destroyed at the end of step $i+1$. It is equivalent to say that no vertex and no edge is in conflict with $s$ over $G(N^{i+1})$. By Lemma 4 (i) it follows that $\text{Vor}_P(s|N^{i+1})$ is empty.

Let now $x \in \text{Vor}_P(s|N^i) \cap \partial \text{Vor}_P(s_{i+1}|N^i)$. Since $x \in \partial \text{Vor}_P(s_{i+1}|N^i)$, there is some $s' \in N^i$ such that $d(x, s_{i+1}) = d(x, s')$. However, since $x \in \text{Vor}_P(s|N^i)$, we have $d(x, s) > d(x, s')$, and so $d(x, s) > d(x, s_{i+1})$. This implies $x \in \text{Vor}_P(s|N^{i+1})$.

A point $p$ is said to be visible from a point $q$ within $G(N^i)$ if there is a path connecting $q$ to $p$ whose interior does not intersect any vertex or edge of $G(N^i)$. Similarly, an edge $e$ is said to be visible from a point $q$ within $G(N^i)$ if there is a point on $e$ that is visible from $q$.

**Fig. 4.** The proof of Lemma 5 (ii). (a) $X(s)$ is a vertex. (b) $X(s)$ is a portion of an edge. In this example, $z = a$ and $x = b$.

**Lemma 6.** Suppose that $s$ is a non-inserted site in $N \setminus N^{i+1}$ and $X(s) \cap \text{Vor}_P(s_{i+1}|N^i) \neq \emptyset$. If $\text{Vor}_P(s|N^{i+1}) \neq \emptyset$, then one of these two cases must occur:

(i) $s$ has a conflicting vertex $v \in \text{Vor}_P(s|N^i) \cap \partial \text{Vor}_P(s_{i+1}|N^i)$ in $G(N^{i+1})$. 


(ii) $S$ has no conflicting vertex in $\text{Vor}_P(s|N^i) \cap \partial \text{Vor}_P(s_{i+1}|N^i)$, but has a single conflicting edge $e \in \partial \text{Vor}_P(s_{i+1}|N^i)$ such that $e$ is visible from a point $p \in G(N^i) \cap \text{Vor}_P(s_{i+1}|N^i) \cap \text{Vor}_P(s|N^i)$ within $G(N^i)$. 

Proof. By Lemma 5, each point of $\text{Vor}_P(s|N^i) \cap \partial \text{Vor}_P(s_{i+1}|N^i)$ is in conflict with $s$. Define $C = \text{Vor}_P(s|N^i) \cap \partial \text{Vor}_P(s_{i+1}|N^i)$. Then $C$ is a connected chain due to Lemma 1. If $C$ contains a vertex of $G(N^{i+1})$, case (i) is true.

If $C$ is a portion of an edge $e$ of $\partial \text{Vor}_P(s_{i+1}|N^i)$, we need to prove that $e$ is visible from some point $p \in G(N^i) \cap \text{Vor}_P(s_{i+1}|N^i) \cap \text{Vor}_P(s|N^i)$. For the following description, see Fig. 4.

Let $z$ be a point of $C$. Let $x$ be a point of $X(s) \cap \text{Vor}_P(s_{i+1}|N^i)$. Note here that if $X(s)$ is a vertex, then $x = X(s)$; if $X(s)$ is an edge, $x$ is any point of $X(s)$ which belongs to $\text{Vor}_P(s_{i+1}|N^i)$. Consider an arbitrary simple path $\gamma$ connecting $x$ to $z$ within $\text{Vor}_P(s|N^i)$ (not within $\text{Vor}_P(s|N^{i+1})$). Since $x$ lies in $\text{Vor}_P(s_{i+1}|N^i)$ and $z$ does not lie in the interior of $\text{Vor}_P(s_{i+1}|N^i)$, $\gamma$ must intersect $\partial \text{Vor}_P(s_{i+1}|N^i)$ at least once, possibly in $z$. We denote by $a$ the first such intersection on the way from $x$ to $z$. Let $e$ be the edge of $\partial \text{Vor}_P(s_{i+1}|N^i)$ containing $a$.

Let $b$ be the last intersection of $G(N^i)$ with $\gamma$ on the way from $x$ to $a$. If $b = x$, then it means $e$ is an edge bounding some region in $G(N^i)$ incident to $x$, so $e$ is visible from $x$, that is, $p = x$. If $b \neq x$, Lemma 4 (iii) implies that an endpoint $p$ of the edge of $G(N^i)$ containing $b$ conflicts with $s$. It means that $a$ is visible from $p$, so $e$ is visible from $p$, and $e$ intersects $\gamma$. Hence $e$ conflicts with $s$ in $G(N^{i+1})$.

Lemma 7. For a site $s \in N \setminus N^i$ and a given vertex or edge of $G(N^i)$, we can decide in $O(\log m)$ time whether they are in conflict.

Proof. A vertex $w \in G(N^i)$ is in conflict with $s$ if and only if $w$ is farther from $s$ than the three sites in $N^i$ that define $w$. This can be checked in constant time because we have access to the coordinates of $w$ and the sites defining the adjacent faces.

Suppose now that we want to check whether an edge $e$ is in conflict with a site $s$. Let $s', s'' \in N^i$ be the sites defining the faces adjacent to $e$, and let $\pi$ be the bisecting plane of $s', s''$. The edge $e$ is embedded in $\pi$, and is in conflict with $s$ if and only if it intersects the halfplane $\pi' := \{x \in \pi \mid d(x, s) > d(x, s')\}$. We test this by ray shooting on $P$ in time $O(\log m)$.

We are now ready to describe how the conflict list is updated during the insertion of a new site $s_{i+1}$. We first collect all not-yet-inserted sites $s$ whose conflict is destroyed by the insertion of $s_{i+1}$. This can be done during the exploration of the subgraph. Since $X(s)$ is being deleted in $G(N^{i+1})$, we need to either find a new conflicting object in $G(N^{i+1})$, or to find out that there is none, which means that $\text{Vor}_P(s|N^j) = \emptyset$ for all $j \geq i + 1$ by Lemma 5.

Suppose that $\text{Vor}_P(s|N^{i+1})$ is not empty. By Lemma 6 we can find a new vertex or edge in conflict with $s$ by exploring the subgraph of $G(N^i)$ belonging
to \(\text{Vor}_P(s|N^i) \cap \text{Vor}_P(s_{i+1}|N^i)\) as follows. If \(X(s)\) is an edge, it may have a conflict with a new vertex of \(\partial\text{Vor}_P(s_{i+1}|N^i)\) that is defined in the interior of \(X(s) \cap \partial\text{Vor}_P(s_{i+1}|N^i)\) or a new edge that is visible from any point in \(X(s) \cap \text{Vor}_P(s_{i+1}|N^i)\). In the latter case, there can be at most two such edges, thus we can check if the edges conflict with \(s\). If they do not, then \(s\) can be discarded since its Voronoi region is empty.

On the other hand, if \(X(s)\) is a vertex, then we first look for a new conflicting vertex by walking among the vertices of \(G(N^i)\) in conflict with \(s\) until we reach a vertex of \(\partial\text{Vor}_P(s_{i+1}|N^i)\). If we do not find such a vertex, we still have to check if there is a new conflicting edge of \(\partial\text{Vor}_P(s_{i+1}|N^i)\) which is visible from some vertex visited so far (see Lemma \(6\)(ii)). If we fail to find a conflicting edge, then \(s\) can be discarded.

6 Analysis

Two of our primitive operations—creating a Voronoi vertex and detecting conflict with an edge—require ray shooting query on \(P\), and therefore take \(O(\log m)\) time. All other primitive operations take constant time. After the initial preprocessing in time \(O(m \log m)\), a backward analysis argument (see pp. 98–108 in [5]) shows that our algorithm performs an expected number of \(O(n \log n)\) primitive operations. This implies a running time of \(O(m \log m + n \log n \log m)\).

**Theorem 2.** Given a polytope \(P\) with \(m\) vertices, and \(n\) points \(N\) on the surface of \(P\), we can compute \(G(N) = G_P(N)\) in expected time \(O(m \log m + n \log n \log m)\).

7 Point Location Queries and Farthest Neighbors

Once we have computed the implicit representation of the farthest-point Voronoi diagram of \(N\) on \(P\), we can easily find a point in the interior of each face \(\text{Vor}_P(s|N)\). For example, take two non-adjacent vertices in an ordinary face, or the two vertices of an eye, and shoot a ray from \(s\) towards their midpoint. Then, run the same algorithm again, but now computing a radial triangulation of \(G(N)\) (see pp. 109 in [5]). Each of its faces is the intersection of \(P\) with a cone \((ss', sv, sv')\) where \(s \in N\), \(s'\) is the interior point for \(\text{Vor}_P(s|N)\), and \(v, v'\) are consecutive vertices of \(\text{Vor}_P(s|N)\). During the second pass, we maintain for each site, whether it has been inserted or not, a pointer to the face of the radial triangulation that contains it. One can see easily that this does not hurt our time bound. It allows to find for each site the face of \(G(N)\) it belongs to. Still following [5] we can build a point location data structure for our radial triangulation that answer queries in \(O(\log^2 n)\) time with high probability. The reason why we need the radial triangulation for these two results is a configuration space argument, namely we want the faces of our graph to have bounded degree.

**Theorem 3.** Given a set \(N\) of points in 3-dimensional convex positions, all-pairs farthest neighbors can be computed in expected \(O(n \log^2 n)\) time. A data
structure that answers farthest-neighbor queries in $O(\log^2 n)$ time with high probability can be build within the same time bounds.

8 Conclusion and Open Problems

The main open problem is to fill the gap between the time complexity of our algorithm, namely $O(n \log^2 n)$, and the lower bound $\Omega(n)$. Besides, one can try to find a deterministic equivalent, but this looks hard. Indeed, a simple $O(n \log n)$ algorithm [4] for the diameter problem has been found in 1989 but a deterministic $O(n \log^2 n)$ [7] time algorithm was only published in 1997.

References

Finding an Optimal Bridge between Two Polygons

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Abstract. Let \( \pi(a, b) \) denote the shortest path between two points \( a, b \) inside a simple polygon \( P \), which totally lies in \( P \). The geodesic distance between \( a \) and \( b \) in \( P \) is defined as the length of \( \pi(a, b) \), denoted by \( gd(a, b) \), in contrast with the Euclidean distance between \( a \) and \( b \), denoted by \( d(a, b) \). Given two disjoint polygons \( P \) and \( Q \) in the plane, the bridge problem asks for a line segment (optimal bridge) that connects a point \( p \) on the boundary of \( P \) and a point \( q \) on the boundary of \( Q \) such that the sum of three distances \( gd(p', p), d(p, q) \) and \( gd(q, q') \), with any \( p' \in P \) and any \( q' \in Q \), is minimized. We present an \( O(n \log^3 n) \) time algorithm for finding an optimal bridge between two simple polygons. This significantly improves upon the previous \( O(n^2) \) time bound.

1 Introduction

The bridge problem, introduced by Cai et al. [2], is defined as: Given two disjoint regions, we want to build a bridge to connect them so as to minimize the length of the longest path between two points in different regions. Because of its similarity to many geometric problems (e.g., diameter problems, minimum separation problems and minimum spanning tree problems), much progress has been made, immediately after the problem is posed. For the simple case where two polygons are convex, two optimal algorithms have been presented by Kim and Shin [5] and by Tan [8], independently. Kim and Shin also reported an \( O(n^2) \) time algorithm for finding an optimal bridge between two simple polygons [5]. A linear time solution to the case of two rectilinear polygons is given in [10]. Efficient algorithms for high-dimensional bridge problems can be found in [8,9].

In this paper, we present an \( O(n \log^3 n) \) time algorithm for finding an optimal bridge between two simple polygons with a total of \( n \) vertices. Our result is obtained by making use of a hierarchical structure that consists of the well-known segment trees, range trees and persistent search trees [7], and a structure that supports dynamic ray shooting and shortest path queries as well [3].

2 Necessary Conditions for Optimal Bridges

Let \( \pi(a, b) \) denote the shortest path between two points \( a, b \) inside a simple polygon \( P \), which starts at \( a \), ends at \( b \) and totally lies in \( P \). The length of
Finding an Optimal Bridge between Two Polygons

\( \pi(a, b) \) is the geodesic distance between \( a \) and \( b \) in \( P \), and is denoted by \( gd(a, b) \). Usually, the Euclidean distance between \( a \) and \( b \) is denoted by \( d(a, b) \).

Let \( P \) and \( Q \) denote two disjoint polygons in the plane with boundaries \( \partial P \) and \( \partial Q \), respectively. The bridge problem considered in [2] deals with finding a line segment that connects a point \( p \in \partial P \) and a point \( q \in \partial Q \) so that the length function \( \max_{p' \in P} \{gd(p', p)\} + d(p, q) + \max_{q' \in Q} \{gd(q, q')\} \) is minimized. We call the segment \((p, q)\) the optimal bridge between \( P \) and \( Q \).

We will first review necessary conditions for optimal bridges between two convex polygons given in [8], and then generalize them to simple polygons.

2.1 The Case of Two Convex Polygons

Notice that in the case of convex polygons, \( gd(p', p) = d(p', p) \) and \( gd(q, q') = d(q, q') \). Let \( V(P) \) denote the furthest-site Voronoi diagram of vertices of polygon \( P \) and \( V(p_i) \) the furthest-site Voronoi region of a vertex \( p_i \) [9]. Let \( p_{i,j} \) denote the intersection point of \( \partial P \) with the common edge of regions \( V(p_i) \) and \( V(p_j) \). Define the infinite triangular region, bounded by two half lines being collinear with but not containing the segments \((p_i, p_{i,j})\) and \((p_j, p_{i,j})\), as the fan of \( p_{i,j} \), and denote it by \( F(p_{i,j}) \) (see Fig. 1b). The fan \( F(p_{i,j}) \) specifies the area to which the candidate bridges with \( p = p_{i,j} \) belong (see Lemma 1). Also, let \( p_{i1} \) and \( p_{i2} \) be two intersection points of \( \partial P \) with the boundary of a region \( V(p_i) \). Define the fan of the vertex \( p_i \), denoted by \( F(p_i) \), as the triangular region bounded by two half lines passing through and containing the segments \((p_i, p_{i1})\) and \((p_i, p_{i2})\). The fan \( F(p_i) \) specifies the area to which the candidate bridges with \( p' = p_i \) belong. The same definition also applies to polygon \( Q \).

Lemma 1 The points \( p \) and \( p' \) of an optimal solution must satisfy one of the following conditions (the same conditions symmetrically hold for the points \( q \) and \( q' \)):

1. \( p \) is the intersection point of \( \partial P \) with the common edge of \( V(p_i) \) and \( V(p_j) \), and the fan \( F(p) \) contains the other point \( q \) of the optimal bridge, or
2. \( p \) is the intersection point of \( \partial P \) with the segment \((p', q)\) and the fan \( F(p') \) contains the point \( q \), where \( q \) satisfies the condition (1), or
3. \( p \) is the intersection point of \( \partial P \) with the segment \((p', q')\), and the fans \( F(p') \) and \( F(q') \) contain the vertices \( q' \) and \( p' \), respectively.

Proof. We will give a simple proof of the condition (1) and omit the proofs of conditions (2) and (3) (see also [8]).

Consider first the case where any two of segments \((p', p)\), \((p, q)\) and \((q, q')\) of an optimal solution are not collinear. Assume that vertex \( p_i \) is the point \( p' \) of the optimal solution. Obviously, a contradiction occurs if \( p_i \) is not the vertex farthest to the point \( p \) of the optimal bridge in polygon \( P \). If \( p_i \) is the only vertex farthest to the point \( p \) in \( P \) (Fig. 1a), or \( p_i \) and another vertex \( p_j \) are the vertices farthest to the point \( p \) (= \( p_{i,j} \)) but the point \( q \) of the optimal bridge is not contained in the fan \( F(p_{i,j}) \) (Fig. 1b), we can move \( p \) by a very small distance to obtain a new point \( p^* \) such that \( p^* \) and \( q \) are mutually visible, \( p_i \) or...
\( p_j \) is still the vertex farthest to \( p^* \), and \( d(p_i, p^*) + d(p^*, q) < d(p_i, p) + d(p, q) \) or \( d(p_j, p^*) + d(p^*, q) < d(p_j, p) + d(p, q) \), a contradiction. Therefore, the points \( p \) and \( p' \) of the optimal solution satisfy the condition (1). \( \Box \)

2.2 The Case of Two Simple Polygons

Given a simple polygon \( P' \) of \( m \) vertices and a set \( S = \{s_1, \ldots, s_k\} \) of \( k \) sites inside \( P' \), the farthest-site geodesic Voronoi diagram of \( S \) in \( P' \) is the partition of \( P' \) into \( k \) regions \( GV(s_1), \ldots, GV(s_k) \) (some of them may be empty) such that a point \( x \in P' \) belongs to \( GV(s_i) \) if and only if the geodesic distance of site \( s_i \) to \( x \) is farther than that of any other site to \( x \). A \( O((m+k) \log(m+k)) \) time algorithm has been presented to compute the farthest-site geodesic Voronoi diagram of \( k \) sites in a simple polygon of \( m \) vertices [1].

Let \( GV(P) \) be the farthest-site geodesic Voronoi diagram of vertices of the simple polygon \( P \). Let \( p_i, p_j \) denote two vertices of \( P \) such that their regions, \( GV(p_i) \) and \( GV(p_j) \), are adjacent, and \( p_{i,j} \) the intersection point of \( \partial P \) with the common edge of \( GV(p_i) \) and \( GV(p_j) \). Similarly, the fan \( F(p_{i,j}) \) needs to be defined. Let \( L_1 \) and \( L_2 \) be two infinite rays shot at \( p_{i,j} \), being collinear with but not containing the first segments of two paths \( \pi(p_{i,j}, p_i) \) and \( \pi(p_{i,j}, p_j) \), respectively, and let \( T(L_1, L_2) \) denote the triangular region bounded by \( L_1 \) and \( L_2 \) (see Fig. 6). Since parts of polygon \( P \) may appear in \( T(L_1, L_2) \), i.e., they block the visibility of the point \( p_{i,j} \), region \( T(L_1, L_2) \) cannot simply be defined as the fan \( F(p_{i,j}) \). Let \( [p^1_{i,j}, p^2_{i,j}] \) be the maximal interval on \( \partial Q \) such that \( p^1_{i,j} \) and \( p^2_{i,j} \) lie in \( T(L_1, L_2) \) and both of them are visible from \( p_{i,j} \). We then define the fan \( F(p_{i,j}) \) as the triangular region bounded by two half lines being collinear with and containing the segments \( (p_{i,j}, p^1_{i,j}) \) and \( (p_{i,j}, p^2_{i,j}) \).

For a non-empty region \( GV(p_i) \), we further divide the boundary of \( GV(p_i) \), overlapped with \( \partial P \), into the intervals such that all points within an interval have (or are visible from) the same first turn point in their shortest paths to \( p_i \). See Fig. 2 for an example. Note that a first turn point may have several intervals. In [1], Aronov et al. referred to such a first turn point as the anchor of all points of the interval with respect to \( p_i \), and defined a refined Voronoi diagram so that all points in a refined region have the same furthest vertex and the same anchor with respect to the furthest vertex.
Let \( p_i(a) \) be an anchor with respect to \( p_i \), and \([p_i(a_1), p_i(a_2)]\) one of its intervals. See Fig. 2. To define the fan \( F(p_i(a)) \) with respect to the interval \([p_i(a_1), p_i(a_2)]\), we modify \( \partial P \) by removing the interval \([p_i(a_1), p_i(a_2)]\), and by inserting the edges \((p_i(a_1), p_i(a_1))\) and \((p_i(a), p_i(a_2))\). The definition of \( F(p_i(a)) \) is then similar to that of \( F(p_1) \), with a slight modification that \( T(L_1, L_2) \) is replaced by the region bounded by two half lines being collinear with and containing \((p_i(a_1), p_i(a_1))\) and \((p_i(a), p_i(a_2))\). (An anchor has as many fans as the intervals it has.) The same definitions also apply to polygon \( Q \).

**Lemma 2** Let \( p'(a) \) and \( q'(a) \) denote the anchors of \( p \) and \( q \) with respect to \( p' \) and \( q' \), respectively. The points \( p \) and \( p' \) of an optimal solution must satisfy one of the following conditions (the same conditions symmetrically hold for the points \( q \) and \( q' \)):

1. \( p \) is the intersection point of \( \partial P \) with the common edge of \( GV(p_i) \) and \( GV(p_j) \), and \( F(p) \) contains the other point \( q \) of the optimal bridge, or
2. \( p \) is the intersection point of \( \partial P \) with the segment \((p'(a), q)\) and \( F(p'(a)) \) contains the point \( q \), where \( q \) satisfies the condition (1), or
3. \( p \) is the intersection point of \( \partial P \) with the segment \((p'(a), q'(a))\), and \( F(p'(a)) \) and \( F(q'(a)) \) contain the anchors \( q'(a) \) and \( p'(a) \) respectively, or
4. \((p,q)\) is a line segment that touches a vertex \( v \) of \( P \) or \( Q \), and three points \( p'(a) \), \( q'(a) \) and \( v \) are in the same side of the line passing through \((p,q)\).

**Proof.** Consider first the case where the optimal bridge \((p,q)\) is not collinear to the first segment of the path \( \pi(p,p') \) or \( \pi(q,q') \). Assume that vertex \( p_i \) is the point \( p' \) of the optimal solution. In the case where \( p_i \) is the only vertex of \( P \) farthest to the point \( p \) in the geodesic metric, or \( p_i \) and another vertex \( p_j \) are the vertices farthest to the point \( p \) (= \( p_{i,j} \)) but the point \( q \) of the optimal bridge is not contained in the fan \( F(p_{i,j}) \), a contradiction occurs, provided that point \( p \) can be moved by a small distance to get a smaller value of \( gd(p', p) + d(p, q) + gd(q, q') \). The condition (1) is satisfied in this case.

The conditions (2) and (3) can be proved by an argument similar to the proof of Lemma 1 (see [S] for a full proof), with a slight modification that \( p \) is the intersection point of \( \partial P \) with the segment \((p'(a), q)\) or \((p'(a), q'(a))\).

Since polygons \( P \) and \( Q \) are simple, the bridge \((p,q)\) may touch a vertex \( v \) of \( P \) or \( Q \). See Fig. 3. (In this case, point \( p \) cannot simply be moved to give a contradiction.) Without loss of generality, assume that \( v \) is a vertex of \( P \). Let \( p_i \)
(resp. $q_j$) denote the vertex of $P$ (resp. $Q$) farthest to the point $p$ (resp. $q$) in the geodesic metric, $p_i(a)$ (resp. $q_i(a)$) the anchor of $p$ (resp. $q$) with respect to $p_i$ (resp. $q_j$), and $q(p_i(a))$ (resp. $p(q_j(a))$) the intersection point of $\partial Q$ (resp. $\partial P$) with the line passing through $(v, p_i(a))$ (resp. $(v, q_j(a))$). See Fig. 3 for an example. Obviously, the optimal bridge cannot be located at the position of $(p^*, q^*)$, where $p^*$ (or $q^*$) is not contained in the triangle determined by $p_i(a), p(q_j(a))$ and $v$; otherwise, $d(p_i(a), p(v)) + d(p, q_j(a)) < d(p_i(a), p^*) + d(p^*, q^*) + d(q^*, q_j(a))$, a contradiction (Fig. 3). Consider a process of rotating the bridge $(p, q)$ around point $v$, with the point $p$ contained in the triangle described above. Assume first that nothing is encountered in the process of rotating the bridge $(p, q)$, i.e., it begins with $p = p(q_j(a))$ and ends with $q = q(p_i(a))$. The sum $d(p_i(a), p) + d(p, v)$ is monotone decreasing in the rotating process, while the sum $d(p, q) + d(p, q) + d(q, q_j(a))$ is monotone increasing. Therefore, the length function $d(p_i(a), p) + d(p, v) + d(v, q) + d(q, q_j(a))$ as well as $gd(p', q) + d(p, q) + gd(q, q')$ is convex with a unique local minimum. If the minimum occurs at point $p^*$ in the rotating process, then $(p^*, q^*)$ is a candidate bridge. This gives the condition (4).

In the rotating process, the bridge $(p, q)$ may touch two vertices of $P$ or/and $Q$, and the vertex $v$ of polygon $P$ or the anchor $q_j(a)$ may be changed. Consider first the situation where the endpoint $q$ of the current bridge reaches the endpoint of $q_j(a)$’s interval on $\partial Q$ (Fig. 4a). In this case, the anchor of $q_j$ is changed to a new one, say, $q_j'(a)$ in Fig. 4a. (The vertex $q_j$ needn’t be changed in the rotating

![Fig. 3. A simple case in the rotation of the bridge $(p, q)$.

![Fig. 4. Two more cases in the rotation of the bridge $(p, q)$.

]
process; otherwise, point \( q \) were not contained in the triangle described above.)

In the situation where the current bridge \((p, q)\) makes a tangent contact with a portion (or two edges) of \( Q \), the candidate bridge should be found before or at the tangent contact (Fig. 4b). In the situation where the current bridge \((p, q)\) touches another vertex \( v' \) of polygon \( P \), the center of the rotation is changed from \( v \) to \( v' \). In all cases, the rotated ranges on both \( \partial Q \) and \( \partial P \) are continuous.

Therefore, the length function considered in the process of rotating the bridge \((p, q)\) is convex, and thus we obtain the condition (4). □

3 The Algorithm

Two types of the fans defined in Section 3.2 are dealt with by the following two lemmas, respectively.

**Lemma 3** Let \( p_i, p_j \) denote two vertices of \( P \) such that \( GV(p_i) \) and \( GV(p_j) \) are adjacent, and \( p_{i,j} \) the intersection point of \( \partial P \) with their common edge. All fans \( F(p_{i,j}) \) for \( p_i, p_j \in P \) can be computed in \( O(n \log^2 n) \) time.

**Proof.** Before describing our algorithm, we briefly review the result of Goodrich and Tamassia on dynamic ray shooting and shortest path queries [3]. For a connected subdivision of the plane into simple polygons (i.e., no “islands” exist), one can maintain a data structure of size \( O(n) \) that supports query time of \( O(\log^2 n) \) for a ray shooting or shortest path in a region of the subdivision (plus \( O(k) \) if the path of length \( k \) is to be reported), and update time of \( O(\log^2 n) \) for each addition/deletion of an edge/vertex.

To use dynamic ray shooting and shortest path algorithms, we construct a planar subdivision that always consists of a bounded region and an unbounded region. The bounded region, denoted by \( PQ \), is an approximation of the exterior of the union \( P \cup Q \). The region \( PQ \) is initially constructed as follows. Let \( a \) be the highest point among all vertices of \( P \) and \( Q \). Place two points \( b \) and \( c \) sufficient below and to the left and right of the union of \( P \) and \( Q \), respectively, and connect three points \( a, b, c \) so that \( P \) and \( Q \) are contained in the triangle \( \triangle abc \). Make a copy \( a' \) of the vertex \( a \). Then, find a point \( p^* \) on \( \partial P \) and a point \( q^* \) on \( \partial Q \) such that they are visible each other. Split each of \( p^* \) and \( q^* \) into two and insert two edges: one connects vertices \( p^* \) and \( q^* \), and the other does their copies. The resulting (simple) polygon \( PQ \) turns the original polygons \( P \) and \( Q \) inside out (Fig. 5).

It is easy to see that a ray shooting query in the exterior of the union \( P \cup Q \) can be performed in \( PQ \) with a slight modification that the ray hits nothing if the hit point is on the boundary of \( \triangle abc \), and an extra ray shooting needs to perform if the hit point is on the pseudo-edge \((p^*, q^*)\) or its copy. Special attention should be paid to the shortest path queries. This is because a shortest path in the exterior of the union \( P \cup Q \) may require to go across the pseudo-edges or even wrap around the highest vertex \( a \). When a shortest path query is performed in \( PQ \), we first find a new pair of mutually visible points, one on \( \partial P \) and the other on \( \partial Q \), such that the shortest path to be computed should not go.
across the line segment connecting them. After the polygon $PQ$ is accordingly modified, the shortest path query can safely be performed. Also, the shortest path between two points $u$ and $v$ in the exterior of $P \cup Q$ is the shortest one of three paths $\pi(u,v)$, $\pi(u,a) \cup \pi(a,v)$ and $\pi(u,a') \cup \pi(a',v)$ in $PQ$.

First, all points $p_{i,j}$ can be obtained from the intersections of $\partial P$ with edges of the diagram $GV(P)$. To compute the interval $[p_{1,i,j},p_{2,i,j}]$ on $\partial Q$, we shoot two rays at $p_{i,j}$ along $L_1$ and $L_2$ in polygon $PQ$. If both of them hit $\partial Q$, the interval $[p_{1,i,j},p_{2,i,j}]$ is obtained and we are done. If only one of them, say, the point hit by the ray along $L_2$, is on $\partial Q$, then we obtain the point $p_{2,i,j}$. To compute the point $p_{1,i,j}$, we modify polygon $PQ$ by inserting the edge connecting $p_{i,j}$ with the hit point on $\partial Q$ and its copy, and by removing the old pseudo-edges. Let $f_2$ denote the intersection point of $\partial Q$ with $L_2$, which is farthest to $p_{i,j}$ along $L_2$. Point $f_2$ can be computed by shooting a ray at infinity in the converse direction of $L_2$ in the exterior of polygon $Q$. See Figs. 6a-b. Then, find the first turn point of the path $\pi(p_{i,j},f_2)$ in the polygon $PQ$. If the turn point is on $\partial Q$, it is just $p_{1,i,j}$ (Fig. 6a). Otherwise, shoot a ray at $p_{i,j}$ to the first turn point (on $\partial P$), and the hit point on $\partial Q$ gives $p_{1,i,j}$ (Fig. 6b).

Consider now the situation where two rays along $L_1$ and $L_2$ hit $\partial P$. Let $p_1$ and $p'_1$ denote two intersection points of $\partial P$ with $L_1$, closest and farthest to $p_{i,j}$, respectively. Modify the polygon $PQ$ by inserting the segment $(p_{i,j},p_1)$ and its copy, and by removing the old pseudo-edges. Then, find the first turn point of the path $\pi(p_{i,j},p'_1)$ in the resulting polygon $PQ$. If the turn point is on $\partial Q$, it is an endpoint of the interval $[p_{1,i,j},p_{2,i,j}]$. Otherwise, shoot a ray at $p_{i,j}$ towards that

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**Fig. 5.** Turning polygons $P$ and $Q$ inside out.

**Fig. 6.** Illustration for the proof of Lemma 3.
turn point. If the hit point is on \( \partial Q \), it gives an endpoint of \([p_{i,j}^1, p_{i,j}^2]\), say, \( p_{i,j}^1 \) in Fig. 6c; otherwise, the fan \( F(p_{i,j}) \) is empty. If \( F(p_{i,j}) \) is not empty, the other endpoint of \([p_{i,j}^1, p_{i,j}^2]\) can similarly be found by considering two intersections of \( \partial P \) with \( L_2 \), one closest to \( p_{i,j} \) and the other farthest to \( p_{i,j} \).

Finally, consider the situation where two rays shot at \( p_{i,j} \) along \( L_1 \) and \( L_2 \) hit nothing. If a point of \( Q \) is contained in \( T(L_1, L_2) \), the convex hull \( CH(Q) \) of \( Q \) is completely contained in \( T(L_1, L_2) \), and thus points \( p_{i,j}^1, p_{i,j}^2 \) can be obtained from two tangent points of \( p_{i,j} \) to \( CH(Q) \). Otherwise, the fan \( F(p_{i,j}) \) is empty.

Since a constant number of operations is needed, the interval \([p_{i,j}^1, p_{i,j}^2]\) can be computed in \( O(\log^2 n) \) time. Hence, the lemma follows. \( \square \)

**Lemma 4** Let \( GV(p_i) \) be a non-empty farthest-site geodesic Voronoi region, and \( A(p_i) \) the set of anchors of all points on the boundary of \( GV(p_i) \) with respect to vertex \( p_i \). All fans \( F(p_i(a)) \), for \( p_i \in P \) and \( p_i(a) \in A(p_i) \), can be computed in \( O(n \log^2 n) \) time.

**Proof.** As it can be proved by an argument similar to the proof of Lemma 3, we omit the detail in this extended abstract. \( \square \)

By now, we can give the main result of this paper using the results obtained in the previous sections.

**Theorem 1** An optimal bridge between two simple polygons \( P \) and \( Q \) can be computed in \( O(n \log^3 n) \) time.

**Proof.** First, compute the furthest-site geodesic Voronoi diagrams of vertices of \( P \) and \( Q \), respectively. As shown in Lemmas 3 and 4, all fans of \( P \) and \( Q \) can be computed in \( O(n \log^2 n) \) time.

It follows from the condition (1) of Lemma 2 that if \( p_i \) and \( p_j \) are two vertices farthest to the point \( p_{i,j} \) in the geodesic metric, then \( p_{i,j} \) is a candidate of the points \( p \) of optimal bridges. In particular, we call such points \( p_{i,j} \) the candidate points. For convenience, we denote by \( p^1, \ldots, p^L \) and \( q^1, \ldots, q^M \) the sequences of candidate points on \( \partial P \) and \( \partial Q \), respectively, and \( w(p^i) \) (\( w(q^j) \)) the geodesic distance from \( p^i \) (\( q^j \)) to the vertex farthest to \( p^i \) in \( P \) (\( q^j \) in \( Q \)).

To quickly find the candidate bridges satisfying the first three conditions of Lemma 2, we construct a hierarchical structure that consists of segment trees, range trees and persistent search trees. For this purpose, we need a new idea to represent the fan of a candidate point. Order all vertices and candidate points on \( \partial P \) (\( \partial Q \)), say, by a clockwise scan of the polygon boundary, starting at a specified vertex \( v_p \in P \) (\( v_q \in Q \)). For a complete ordering, we consider the point \( v_p \) as two points \( v_{pl} \) and \( v_{pr} \) such that \( v_{pl} \leq p \leq v_{pr} \), for all points \( p \) on \( \partial P \). The treatment for the point \( v_q \) is the same. Let \( p^{i1} \), \( p^{i2} \) denote the first intersection points of \( \partial Q \) with two bounding half lines of \( F(p^i) \). The fan \( F(p^i) \) can then be represented by the interval \([p^{i1}, p^{i2}]\) if \( v_q \) is not contained in \( F(p^i) \), or two intervals \([v_{ql}, p^{i1}]\) and \([p^{i2}, v_{qr}]\) otherwise.

Consider now how to compute an optimal bridge \((p, q)\) that is not collinear to the first segment of the path \( \pi(p, p') \) or \( \pi(q, q') \). For each point \( p^i \) (\( 1 \leq i \leq l \)), we want to find a point \( q^j \) such that \( p^i \) and \( q^j \) lie in \( F(q^j) \) and \( F(p^i) \), respectively,
and $d(p^i, q^i) + w(q^i)$ is minimized. (In this case, $w(p^i) + d(p^i, q^i) + w(q^i)$ is also minimized.) First, organize the intervals of all fans $F(q^i)$ ($1 \leq q^i \leq m$) into a segment tree. An interval is stored in $O(\log n)$ nodes, and thus each node $s$ of the segment tree has an associated set $COVER(s)$ of the intervals. For each $COVER(s)$, store the candidate points whose corresponding intervals are in $COVER(s)$ into a second-level range tree. Again, a candidate point is stored in $O(\log n)$ nodes, and each node $r$ of the range tree for $COVER(s)$ has an associated set $SUB(r)$ of candidate points. For each $SUB(r)$, we further compute the additively weighted (nearest) Voronoi diagram with the weights $w(q^i)$ for all candidate points $q^j$ in $SUB(r)$, and organize the diagram into a third-level persistent search tree. The resulting data structure is called the segment-range-persistent tree. Since it takes $O(n \log n)$ time and space to construct a segment or range tree, and $O(n \log n)$ time and $O(n)$ space to compute a Voronoi diagram and a persistent search tree, the segment-range-persistent tree can be constructed in $O(n \log^3 n)$ time and $O(n \log^2 n)$ space.

Given a query point $p^i$, we find the point $q^j$ which is optimal with respect to $p^i$ as follows. In the top-level segment tree, find the sets $COVER(s)$ whose intervals contain $p^i$. In the second-level range trees storing the found sets $COVER(s)$, we further find the sets $SUB(r)$ whose candidate points (on $\partial Q$) are contained in the interval $[p^{i1}, p^{i2}]$ or in two intervals $[v_{ql}, p^{i1}]$ and $[p^{i2}, v_{qr}]$. In the third-level persistent search trees, find the candidate points of $Q$ which are locally nearest to $p^i$ using the point location algorithm. The point $q^j$ optimal with respect to $p^i$ is then found from answers of these point location queries. Clearly, the query time taken for $p^i$ is $O(\log^3 n)$. Since the number of candidate points on $\partial P$ is $O(n)$, the time taken to enumerate all of such candidate bridges is $O(n \log^3 n)$.

Analogously, the candidate bridges satisfying the conditions (2) and 3 of Lemma 2 can be computed in $O(n \log^3 n)$ time. We omit the detail in this extended abstract.

To find the candidate bridges satisfying the condition (4), we consider the bounding half lines of anchor’s fans which touch a vertex of $P$ or $Q$. Assume that a bounding line of the fan $F(p_i(a))$ touches a vertex $v$ of $P$ and that $q(p_i(a))$ is the first intersection point of the bounding line with $\partial Q$. Let $q_j$ denote the vertex of $Q$ furthest to $q(p_i(a))$ in the geodesic metric, and $q_j(a)$ the anchor of $q(p_i(a))$ with respect to $q_j$. If $q_j(a)$ and $v$ are in the same side of the line through $(p_i(a), q(p_i(a)))$, we should perform the rotating process described in the proof of Lemma 2.

Assume that $p_1$ is the intersection point of $\partial P$ with the segment $(p_i(a), v)$, and $p_2$ the endpoint of the edge of $P$ which contains $p_1$ and is contained in $F(p_i(a))$. Consider the process of rotating the bridge $(p, q)$, with point $p$ contained in the edge $(p_1, p_2)$. First, we discuss how to handle the events (places) where rotation centers are changed. To find the first place where the center is changed, we compute the first turn point of the path $\pi(v, q_j(a))$ in the exterior of $P$, and the intersection point of the edge $(p_1, p_2)$ with the line passing through $v$ and the first turn point. (The first turn point of $\pi(p_1, q_j(a))$ in the exterior of $P$ can be found in logarithmic time.) If the intersection point does not exist,
the center is kept unchanged in the process of rotating bridge \((p, q)\), with point \(p\) contained in edge \((p_1, p_2)\). If the intersection point is found, the rotation center is changed from \(v\) to that turn point when \(p\) reaches the intersection point. In this case, we consider \((p_1, p_2)\) as two edges, separated by that intersection point. This allows us to consider only the situation where the center is kept unchanged in the process of rotating bridge \((p, q)\), with point \(p\) contained in edge \((p_1, p_2)\).

To find the candidate bridge quickly, we perform a binary search on the edges of \(Q\) which are rotated, using the distance computed for edge \((p_1, p_2)\) and the chosen edge of \(\partial Q\). (Given \(p_i(a), v, q_j(a)\) and a pair of the edges of \(P\) and \(Q\), the position of \(p\) and \(q\) on the edges giving the local minimum can be computed in constant time.) Then, what is the range of the rotated edges on \(\partial Q\)? Clearly, the point \(q(p_i(a))\) is the starting point of the rotated range. However, it takes \(O(n)\) time to exactly give the ending point, as it may require to follow a sequence of anchors \(q_j(a)\). To overcome this difficulty, we temporarily consider it as the endpoint of \(GV(q_j)\) towards which the point \(q\) moves. (Recall that point \(q_j\) cannot be changed in the rotating process.) We then perform the binary search on the edges of \(Q\) contained in the range described above. Whenever the points \(p, q\) giving the local minimum do not exist in two considered edges, they occur like the segment \((p^*, q^*)\) shown in Fig. 3, or the point of \(\partial Q\) hit by the ray shot at \(v\) towards \(q\) in the exterior of \(Q\) is not \(q\) (i.e., the segment \((v, q)\) intersects with \(\partial Q\)), we consider the computed distance as \(\infty\). This makes sure that the binary search should report a pair of points \(p\) and \(q\) giving the local minimum, with the point \(q\) contained in the real range of the rotated edges on \(\partial Q\). If the reported point \(p\) is not identical to \(p_2\), then \((p, q)\) gives a candidate bridge. Otherwise, we need to perform the same procedure for the edge of \(\partial P\), which is next to \((p_1, p_2)\) and contained in \(F(p_i(a))\). For this edge, the starting endpoint of the rotated range on \(\partial Q\) is already known. In this way, we can find all candidate bridges satisfying the condition (4) in \(O(n \log n)\) time.

Finally, we simply report the one among all found bridges which minimizes the maximum distance between \(P\) and \(Q\). It completes the proof. \(\square\)

References

How Good Is Sink Insertion?

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Abstract. Generating high quality meshes is one of the most important steps in many applications such as scientific computing. Sink insertion method is one of the mesh quality improvement methods that had been proposed recently. However, it is unknown whether this method is competitive to generate meshes with small number of elements. In this paper, we show that, given a two-dimensional polygonal domain with no small angles, the sink insertion method generates a well-shaped mesh with $O(n)$ triangles, where $n$ is the minimum number of triangles generated by any method with the same quality guarantee. We also show that the sink insertion method more likely can not guarantee the same result for a three-dimensional domain, while the other methods such as Delaunay refinement can achieve.

Keywords: Mesh generation, Delaunay triangulations, sink insertion, computational geometry, algorithms.

1 Introduction

Mesh generation is the process of breaking a geometric domain into a collection of primitive elements such as triangles in 2D and tetrahedra in 3D. It has plenty of applications in scientific computing, computer graphics, computer vision, geometric information system, and medical imaging. Some applications have a strict quality requirement on the underlying meshes used. For example, most of the numerical simulations require that the mesh is well-shaped. In addition, some numerical simulations methods, for example, the control volume method, prefer the mesh to be a Delaunay triangulation.

Recently, Edelsbrunner et. al [23] proposed a new mesh improvement method based on sink insertion. This new method guarantees to generate a Delaunay mesh with a small radius-edge ratio. Edelsbrunner and Guoy (private communication) found that the sink insertion method tends to be more economical when we want to add as many points as possible at the same time to refine the mesh while maintaining the Delaunay property. It will also be useful in the software environment with off-line Delaunay triangulation or parallel meshing. In stead of dealing with all the circumcenters as many as the number of bad elements, they deal with small number of sinks of these bad elements. From experiments, they observe as many as 100 bad tetrahedra sharing the same sink. However, unlike Delaunay refinement, it is an open problem whether the sink insertion method generates an almost-good Delaunay mesh with $O(n)$ simplex elements,
where \( n \) is the minimum number of \( d \)-dimensional simplex elements generated by any other methods with the same radius-edge ratio quality guarantee.

In this paper, we show that the sink insertion method guarantees to generate a well-shaped mesh with size \( O(n) \) in 2D. For a three-dimensional domain, unlike the Delaunay refinement methods, the size optimality is not guaranteed because of the existence of slivers in an almost-good Delaunay mesh. We give an example that suggests that the sink insertion method may not guarantee to generate a mesh with size \( O(n) \) for a three-dimensional domain.

The rest of the paper is organized as follows. In Section \([2]\) we review the sink insertion algorithm proposed by Edelsbrunner \( et. \ al \). In Section \([3]\) we prove that the sink insertion method guarantees to generate a well-shaped mesh with size \( O(n) \) for a two-dimensional PLC domain with no small angles. In section \([4]\) we discuss why the sink insertion method may not guarantee that the generated mesh has size \( O(n) \) for a three-dimensional PLC domain. Section \([5]\) concludes the paper with further discussions.

### 2 Preliminary

A simplicial mesh is called \textit{almost good} if each of its simplex elements has a small radius-edge ratio, which is the circumradius divided by the shortest edge length of the simplex. Hereafter we use \( \rho(\tau) \) to denote the radius-edge ratio of a simplex \( \tau \). Several theoretical and practical approaches have been proposed for generating almost-good Delaunay meshes. Assume the spatial domain that does not have small angles is given in terms of its piecewise linear complex boundary (\( PLC \)) \([7]\). It has been shown that the Delaunay refinement methods \([1,5,6]\) generate an almost-good Delaunay mesh with size \( O(n) \), where \( n \) is the minimum number of elements for any mesh with the same radius-edge ratio for the same geometric domain.

#### 2.1 Sink

Edelsbrunner \( et. \ al. \) \([2,3]\) defined the sink of a simplex \( \sigma \) in a Delaunay complex by the following recursive approach.

**Definition 1.** [SINK] In a \( d \)-dimensional Delaunay complex, let \( c_\sigma \) be the circumcenter of a \( d \)-simplex \( \sigma \); let \( \mathcal{N}(\sigma) \) be the set of \( d \)-simplices that share a \( d-1 \) dimensional face with \( \sigma \). For each simplex \( \tau \in \mathcal{N}(\sigma) \), let \( H_\tau \) be the half space containing \( \tau \) bounded by the \( d-1 \) dimensional face shared by \( \tau \) and \( \sigma \). A point \( z \) is a sink of \( \sigma \) when

- \( z \) is \( c_\sigma \) and it is contained in \( \sigma \); or
- \( z \) is a sink of \( \tau \in \mathcal{N}(\sigma) \) and \( c_\sigma \) is contained in \( H_\tau \).

A simplex containing its own circumcenter is called a \textit{sink simplex}. Edelsbrunner \( et. \ al. \) \([2,3]\) showed that there is no loop in the definition of sink among all \( d \)-dimensional simplices by proving that the circumradius of the simplex \( \tau \)
containing the sink of a simplex $\sigma$ is not less than the circumradius of $\sigma$. Notice that if the circumcenter of a boundary simplex is not inside the domain, then its sink is not defined by the above definition. For this case, we just define its circumcenter as its sink. Given a boundary $k$-simplex $\sigma$ ($k < d$) contained in a $k$-dimensional boundary polyhedron, its sink is defined by considering only that boundary polyhedron.

The min-circumsphere of a $k$-simplex $\tau$ in $d$-dimensions is the smallest $d$-dimensional sphere that contains all vertices of $\tau$ on its surface. When $k = d$, the min-circumsphere is also called the circumsphere. A point is said to encroach the domain boundary if it is contained inside the min-circumsphere of a boundary $k$-simplex, where $k < d$.

Let $T$ be the set of tetrahedron in a 3-dimensional Delaunay mesh and $T = T \cup \{\tau_\infty\}$, where $\tau_\infty$ represents the outside of the domain, called dummy element. Edelsbrunner et al. [2] defined the flow relation $\prec \subseteq T \times T$ with $\tau_1 \prec \tau_2$ if

1. $\tau_1$ and $\tau_2$ share a common triangle $\nu$, and
2. the interior of $\tau_1$ and the circumcenter $c$ of $\tau_1$ lie on the different side of the plane containing $\nu$.

If $\tau_1 \prec \tau_2$, then $\tau_1$ is called the predecessor of $\tau_2$; and $\tau_2$ is called the successor of $\tau_1$. Here predecessor and successor are only meaningful for a Delaunay tetrahedron. The set of descendants of tetrahedron $\tau$ is defined as

$$\text{desc}(\tau) = \{\tau\} \cup \text{desc}_{\tau \prec \mu}(\mu), \quad \text{where } \text{desc}_{\tau \prec \mu}(\tau) = \bigcup_{\tau \prec \mu} \text{desc}(\mu).$$

Notice that for a triangle, there is only one successor defined, while there are only at most two successors defined for a tetrahedron. A sequence of tetrahedra with $\tau_1 \prec \tau_2 \ldots \prec \tau_n$ is called a flow path from $\tau_1$ to $\tau_n$, denoted by $\pi(\tau_1, \tau_n)$. See the left figure in Figure 1 for an illustration in 2D.

### 2.2 Sink Insertion Algorithms

Sink insertion method, proposed by Edelsbrunner et al. [2], inserts the sinks of bad $d$-dimensional simplex elements instead of inserting their circumcenters directly. A simplex element is bad if its radius-edge ratio is larger than a constant $\varrho$. For the completeness of the presentation, we review the sink insertion algorithm for a three-dimensional domain.

**Algorithm:** Sink-Insertion($\varrho_0$)

**Empty Encroachment:** For any encroached boundary segment, add its midpoint and update the triangulation. For any encroached boundary triangle, add its sink and update the triangulation. If the sink to-be-added encroaches any boundary segment, we split that segment instead of adding that sink.

**Bad Elements:** For any tetrahedron $\sigma$ with $\rho(\sigma) > \varrho_0$, find its sink $s_\sigma$. Assume that $s_\sigma$ is the circumcenter of a tetrahedron $\tau$. Insert the sink $s_\sigma$ to split $\tau$ and update the Delaunay triangulation. However, if $s_\sigma$ encroaches a boundary segment or triangle, we apply the following rules instead of adding $s_\sigma$. 
Equatorial Sphere: For any boundary triangle $\mu$ encroached by the sink $s_\sigma$, add the sink $s_\mu$ of $\mu$. Update the triangulation accordingly. However, if $s_\mu$ encroaches any boundary segment, we apply the following rule instead.

Diametral Sphere: For any boundary segment $\nu$ encroached by the sink $s_\sigma$ or the sink $s_\mu$, add the midpoint of $\nu$. Update the triangulation accordingly.

Recall that the insertion of the circumcenter of a bad $d$-simplex will immediately remove the simplex. Inserting the sink of a bad tetrahedron may seem counter-intuitive, because the sink of a $d$-simplex could be far away from it. Consequently, the insertion of the sink may not remove the bad $d$-simplex immediately. The termination of the algorithm may be in jeopardy. However, it is proved that the circumradius of a tetrahedron $\sigma$ is no more than that of the tetrahedron $\tau$ containing the sink of $\sigma$. Then the proofs of the termination of Delaunay refinement method can be applied directly to prove the termination of the sink insertion algorithm. If we select $\varrho_0 > 1$, then the minimum distance among mesh vertices after the sink insertion will not decrease, which implies the algorithm’s termination. If there are boundary constraints, the constant $\varrho_0$ has to be increased to $\sqrt{2}$ for 2D domain and 2 for 3D domain.

3 Good Grading Guarantee in 2D

This section is devoted to study the number of elements in the generated two-dimensional mesh by analyzing the relation between the nearest neighbor function $N()$ defined by the final mesh and the local feature size function $lfs()$ defined by the input domain. Here $N(x)$ is the distance from $x$ to its second nearest mesh vertex. A mesh vertex $v$ is always the nearest mesh vertex of itself. Local feature size $lfs(x)$ is the radius of the smallest disk centered at $x$ intersecting two non-incident input segments or input vertices. Both $N()$ and $lfs()$ are 1-Lipschitz function. A mesh is said to have a **good grading** if the nearest neighbor function $N()$ defined on the mesh is within a constant factor of $lfs()$.

We study the spacing relations among intermediate meshes by using similar idea as Ruppert and Shewchuk did. With each vertex $v$, we associate an *insertion edge length* $e_v$ equal to the length of the shortest edge connected to $v$ immediately after $v$ is introduced into the Delaunay mesh. Here $v$ may not have to be inserted into the mesh actually. For the sake of convenience of analyzing, we also define a *parent vertex* $p(v)$ for each vertex $v$, unless $v$ is an input vertex. Intuitively, for any non-input vertex $v$, $p(v)$ is the vertex “responsible” for the insertion of $v$. We discuss in detail what means by responsible here. If $v$ is inserted as the sink of a triangle $\sigma$ with $\rho(\sigma) \geq \varrho_0$, then $p(v)$ is the most recently inserted end point of the shortest edge of $\sigma$. If all vertices of $\sigma$ are original input vertices, then $p(v)$ is one of the end points of the shortest edge of $\sigma$. If $v$ is the midpoint of an encroached segment, then $p(v)$ is the encroaching vertex. For the sake of simplicity, we always assume that the encroaching vertex is not an input vertex, because Ruppert and Shewchuk showed that the nearest neighbor function $N()$ defined on the Delaunay mesh after enforcing the domain boundary is within
a constant factor of the local feature size function, i.e., \( N(v) \sim lfs(v) \). The parent vertex \( p(v) \) of \( v \) does not need to be inserted into the mesh actually.

We then show that \( e_v \) of any introduced mesh vertex \( v \) is related to that of its parent vertex \( p(v) \). Here \( v \) may also not be inserted due to encroaching. For a vertex \( v \), as \([6,4]\), we define the density ratio at point \( v \) as \( D_v = \frac{lfs(v)}{e_v} \). Clearly, \( D_v \) is at most one for an input vertex \( v \), and for newly inserted vertex \( v \), \( D_v \) tends to become larger. Notice that \( D_v \) is defined just immediately after \( v \) is introduced to the mesh; it is not defined based on the final mesh.

**Lemma 1. [Radius-variation]** Consider a triangle \( \sigma \). Let \( u \) and \( e_u \) be the circumcenter and the circumradius of \( \sigma \). Assume that there is no triangle with a large radius-edge ratio in \( \text{desc}(\sigma) \) except possibly \( \sigma \) itself. Assume triangle \( \tau \neq \sigma \) contains the sink of \( \sigma \) inside. Let \( \mu \) be the predecessor of triangle \( \tau \). Let \( w \) and \( e_w \) be the circumcenter and circumradius of the triangle \( \mu \). Then

\[
e_w - e_u \geq \frac{1}{4\rho_0} ||w - u||.
\]

**Proof.** Let’s consider a triangle \( \tau_1 = pqr \) and its successor \( \tau_2 = pqs \) in \( \text{desc}(\sigma) \), where \( \tau_2 \neq \tau \). Triangle \( \tau_2 \) has a small radius-edge ratio from the assumption of \( \text{desc}(\sigma) \). Let \( x \) and \( y \) be the circumcenter of \( \tau_1 \) and \( \tau_2 \) respectively. Let \( e_x \) and \( e_y \) be the circumradius of \( \tau_1 \) and \( \tau_2 \) respectively. The right figure in Figure 1 illustrates the proof that follows. Let \( c \) be the midpoint of the edge \( pq \). Let \( \alpha = \angle cqx \) and \( \beta = \angle xqy \). Assume that \( cq \) has a unit length. Then \( e_x = 1/\cos(\alpha) \), \( e_y = 1/\cos(\alpha + \beta) \), and \( ||x - y|| = \tan(\alpha + \beta) - \tan(\alpha) \). It is easy to show that

\[
e_y - e_x = \frac{\sin(\alpha + \beta)}{\cos(\frac{\beta}{2})} = \sin(\alpha + \beta) - \cos(\alpha + \beta) \tan(\frac{\beta}{2}).
\]

Assume we fix the value of \( \alpha + \beta = \theta_0 \), then it is easy to show that

\[
\frac{\sin(\alpha + \beta)}{\cos(\frac{\beta}{2})} = \sin(\theta_0) - \cos(\theta_0) \tan(\frac{\theta_0}{2}) = \tan(\frac{\theta_0}{2}).
\]
Let $ps'$ be a diameter of the circumcircle of the triangle $pq$. It is easy to show that $||q - s|| < ||q - s'|| = 2||c - y|| = 2e_y \sin(\alpha + \beta)$. The triangle $\tau_2 = pqs$ has a small radius-edge ratio implies that $||q - s|| \geq \frac{e_y}{\rho_0}$. Thus we have $\sin(\alpha + \beta) \geq \frac{1}{2\rho_0}$. It follows that $\frac{e_y - e_x}{||q - y||} \geq \tan(\frac{\theta_0}{2}) = \frac{\sin(\theta_0)}{1 + \cos(\theta_0)} > \frac{1}{4\rho_0}$. Using the triangle inequality, it is easy to show that $e_w - e_u > \frac{1}{4\rho_0}||w - u||$ by summing up the inequalities for all triangle and successor pairs $\tau_1$ and $\tau_2$ in desc($\sigma$) with $\tau_2 \neq \tau$.

**Theorem 1.** [Bounded Density Ratio] There are fixed constants $D_1 \geq 1$ and $D_2 \geq 1$ such that for any vertex $v$ inserted or rejected as the sink of a bad triangle, $D_v \leq D_2$; for any vertex $v$ inserted or rejected as the midpoint of an encroached boundary segment, $D_v \leq D_1$. Hence, there is a constant $D = \max\{D_1, D_2\}$ such that $D_v \leq D$ for all mesh vertex $v$.

**Proof.** We prove this by induction. If $v$ is an original input vertex, then the length $e_v$ of the shortest edge connected to $v$ is at least $lfs(v)$ from the definition of $lfs(v)$. Thus $D_v = \frac{lfs(v)}{e_v} \leq 1$ and the theorem holds.

Then consider non-input vertex $v$. We first consider inserting $v$ as the sink of a triangle $\sigma$. Let $u$ be the circumcenter of the triangle $\sigma$. Notice that $v$ is also the sink of any triangle from desc($\sigma$). Without loss of generality, assume that no triangle except $\sigma$ from desc($\sigma$) has a large radius-edge ratio. Assume $v$ is the circumcenter of a triangle $\tau$. Notice that $e_v$ is equal to the circumradius of $\tau$.

Case 1: the triangles $\sigma$ and $\tau$ are the same. Notice that $\sigma$ has a radius-edge ratio at least $\rho_0$, then parent $p$ of vertex $v$ is one of the end points of the shortest edge $pq$ of $\sigma$. Here $p$ could be the most recently inserted vertex or an original vertex of $\sigma$. Then $q$ is an original vertex or is inserted before $p$. In both cases, we have $e_p \leq ||p - q||$. Then $e_p \leq ||p - q|| \leq \frac{R_p}{\rho_0}$. And $e_u = R_\sigma \geq \rho_0 \cdot e_p$. Notice that $lfs(p) \leq D_pe_p$, where $D_p$ is the density ratio bound of vertex $p$ derived from induction. Thus $lfs(u) \leq lfs(p) + ||p - u|| \leq D_pe_p + e_u \leq \left(\frac{D_p}{\rho_0} + 1\right)e_u$. It implies that $D_u = \frac{lfs(u)}{e_u} \leq \frac{D_p}{\rho_0} + 1$. So a sufficient condition for bounding the density ratio of vertex $u$ is

$$\max(D_1, D_2) \frac{1}{\rho_0} + 1 \leq D_2 \tag{1}$$

Case 2: the triangles $\sigma$ and $\tau$ are not the same. Let $w$ be the circumcenter of the predecessor triangle of $\tau$. Similarly we define $e_w$ as the circumradius of that triangle. Then by previous lemma $\square$, we know that there is a constant $\delta = \frac{1}{4\rho_0}$ such that $e_w - e_u \geq \delta||w - u||$. Here $w$ and $u$ could be the same.

Subcase 2.1: vertices $w$ and $u$ are not close, i.e., $||w - u|| \geq \epsilon e_v$, where $\epsilon = \frac{1}{2\rho_0}$. Then similar to the previous lemma $\square$ we have $lfs(v) \leq lfs(u) + ||u - v|| \leq lfs(u) + \frac{1}{2}(e_v - e_u) \leq (1 + \frac{D_p}{\rho_0} - \frac{1}{2})e_u + \frac{1}{2}e_v$. Thus a sufficient condition that $D_v = \frac{lfs(v)}{e_v} \leq D_2$ is $(1 + \frac{D_p}{\rho_0} - \frac{1}{2})e_u \leq (D_2 - \frac{1}{2})e_v$. From $e_u \leq e_v$, this inequality is satisfied if $1 + \frac{D_p}{\rho_0} - \frac{1}{2} \leq D_2 - \frac{1}{2}$, and $D_2 - \frac{1}{2} \geq 0$. From $D_p \leq \max(D_1, D_2)$, a sufficient condition that $D_v \leq D_2$ is

$$D_2 \geq \max\left(\frac{1}{2}, 1 + \frac{\max(D_1, D_2)}{\rho_0}\right). \tag{2}$$
Subcase 2.2: vertices \( w \) and \( v \) are close, i.e., \( ||v-w|| < \epsilon e_v \), where \( \epsilon = \frac{1}{2\varrho_0} \). For vertex \( w \), similar to subcase 2.1, we have \( lfs(w) \leq (1 + \frac{D_v}{\varrho_0})e_u + \frac{1}{\delta}e_w \). Then \( lfs(v) \leq lfs(w) + ||v-w|| \) implies that \( lfs(v) \leq (1 + \frac{D_v}{\varrho_0} - \frac{1}{\delta})e_u + \frac{1}{\delta}e_w + \epsilon e_v \leq (1 + \frac{D_v}{\varrho_0} - \frac{1}{\delta})e_u + (1 + \epsilon)e_w \). Thus, from \( e_u \leq e_v \), a sufficient condition that \( D_v \leq D_2 \) is \( 1 + \frac{D_v}{\varrho_0} - \frac{1}{\delta} \leq D_2 - \frac{1}{\delta} - \epsilon \), and \( D_2 - \frac{1}{\delta} - \epsilon \geq 0 \). Consequently, we need
\[
D_2 \geq \max(1 + \epsilon + \max(D_1, D_2), \frac{1}{\delta} + \epsilon) \quad (3)
\]

Case 3: vertex \( v \) is the midpoint of a segment that is encroached by a vertex \( w \). Here \( w \) is the sink of a triangle \( \sigma \) with large radius-edge ratio. Assume that the sink \( w \) is contained in triangle \( \tau \). Let \( u \) be the circumcenter of triangle \( \sigma \).

Subcase 3.1: vertices \( w \) and \( u \) are not same. From the analysis of case 2, we have \( lfs(w) \leq (1 + \frac{D_v}{\varrho_0} - \frac{1}{\delta})e_u + (1 + \epsilon)e_w \) by substituting \( v \) by \( w \) in the results. The vertex \( w \) is inside the circumcircle centered at \( v \) with radius \( e_v \). Therefore \( e_w \leq \sqrt{2}e_v \). From \( lfs(v) \leq lfs(w) + ||w-u|| \leq (1 + \frac{D_v}{\varrho_0} - \frac{1}{\delta})e_u + (\sqrt{2}\varrho + \sqrt{2}\epsilon + 1)e_v \), inequality \( D_v \leq D_1 \) holds if \( 1 + \frac{D_v}{\varrho_0} - \frac{1}{\delta} \leq D_1 - \sqrt{2}(\frac{1}{\delta} + \epsilon) - 1 \) and \( D_1 \geq \sqrt{2}(\frac{1}{\delta} + \epsilon) + 1 \). Consequently, a sufficient condition is that
\[
D_1 \geq \max\{2 + \sqrt{2}(\frac{1}{\delta} + \epsilon) - \frac{1}{\delta} + \max(D_1, D_2), \sqrt{2}(\frac{1}{\delta} + \epsilon) + 1\} \quad (4)
\]

Subcase 3.2: vertices \( u \) and \( w \) are the same. We have \( lfs(v) \leq lfs(u) + ||u-v|| \leq (1 + \frac{D_v}{\varrho_0})e_u + e_v \leq (1 + \frac{D_v}{\varrho_0})\sqrt{2}e_v + e_v \). To prove \( D_v \leq D_1 \), we need
\[
\sqrt{2}(1 + \frac{D_v}{\varrho_0}) + 1 \leq D_1 \quad (5)
\]

As conclusion, if we choose \( D_1 \) and \( D_2 \) as the follows,
\[
D_1 = \max\left\{ \frac{1 + \sqrt{2}(\frac{1}{\delta} + \epsilon)}{\varrho_0(2\delta + \sqrt{2} - 1 + \sqrt{2}\epsilon)}, \frac{(\sqrt{2} + 1)e_v}{\varrho_0 - \sqrt{2}}, \frac{e_v - \sqrt{2}}{\varrho_0 - \sqrt{2}} \right\}, \quad \text{and} \quad D_2 = \max\left( \frac{1}{\delta} + \epsilon, 1 + \epsilon + \frac{D_1}{\varrho_0} \right),
\]
then all inequalities are satisfied. Notice here \( \delta = \frac{1}{4\varrho_0} \), and \( \epsilon = \frac{1}{2\varrho_0} \).

The following theorem concludes that the generated mesh has a good grading, i.e., for any mesh vertex \( v \), \( N(v) \) is at least some constant factor of \( lfs(v) \).

**Theorem 2.** [Good Grading] For any mesh vertex \( v \) generated by the sink insertion method, the edge incident on \( v \) has length at least \( \frac{lfs(v)}{D+1} \).

The proof is omitted here. The values corresponding to \( D_1 \) and \( D_2 \) guaranteed by the Delaunay refinement method \([5][6]\) are small: \( D_v = \frac{lfs(v)}{e_v} \) is at least \( \frac{1}{\varrho_0 - \sqrt{2}} \) for a vertex \( v \) inserted as the circumcenter of a bad triangle and at least \( \frac{\varrho_0}{\varrho_0 - \sqrt{2}} \).
for a vertex \( v \) inserted as the midpoint of an encroached segment. For instance, Ruppert claims that if the smallest angle is 10 degrees, then no edge is smaller than \( \frac{1}{6} \) of the local feature size of either of its endpoints. To guarantee the minimum angle 10 degrees, we need \( \varrho_0 = \frac{1}{2 \sin 10^\circ} \approx 2.88 \). Then \( \delta \approx 0.087 \) and \( \epsilon \approx 0.174 \). So \( D_1 \approx 17.54 \) and \( D_2 \approx 16.69 \). It then implies that no edge is smaller than \( \frac{1}{19} \) of the local feature size of either of its endpoints in any mesh generated by the sink insertion method. Therefore, the theoretical bound on the number of elements of the mesh generated by sink insertion method is more likely larger than that by the Delaunay refinement method.

Ruppert shows that the nearest neighbor value \( N(v) \) of a mesh vertex \( v \) of any almost-good mesh is at most a constant factor of \( lfs(v) \), where the constant depends on the radius-edge ratio. The above Theorem 2 shows that the nearest neighbor \( N(v) \) for the 2D Delaunay mesh generated by sink insertion is at least some constant factor of \( lfs(v) \). Then we have the following theorem.

**Theorem 3.** [Linear Size] The number of triangles in the 2D mesh generated by the sink insertion method is within a constant factor of any Delaunay mesh for the same domain, where the constant depends on the radius-edge ratios of the meshes.

4 Discussions for 3D Domain

Shewchuk [6] showed that the Delaunay refinement method generates almost-good meshes with a good grading guarantee in two and three dimensions. We had showed that the sink insertion method also generates a almost-good mesh with a good grading guarantee in two dimensions. Unfortunately, the proofs presented here can not be directly applied to three dimensions. The reason is as follows. To guarantee the size optimality of the sink insertion method, the nearest neighbor function \( N() \) defined on the generated mesh must be within a constant factor of the local feature size function \( lfs() \). Notice that \( lfs(v) \geq e_v \geq N(v) \), where \( e_v \) is the length of the shortest edge connected to \( v \) when vertex \( v \) is inserted. Therefore, when a vertex \( v \) is introduced to the mesh, \( e_v \) should be within a constant factor of \( lfs(v) \) to guarantee the good grading of the generated mesh. In other words, we have to prove the existence of a constant \( D \) such that for each vertex \( v \) inserted into the mesh, \( lfs(v) \leq De_v \). Or more specifically, there should exist three constants \( D_1, D_2, \) and \( D_3 \) such that

\[ - \text{ for each vertex } v \text{ inserted as the circumcenter of a tetrahedron, } lfs(v) \leq D_3e_v; \]
\[ - \text{ for } v \text{ inserted as the circumcenter of a boundary triangle, } lfs(v) \leq D_2e_v; \]
\[ - \text{ for } v \text{ inserted as the midpoint of a boundary segment, } lfs(v) \leq D_1e_v. \]

Assume that we insert a vertex \( v \) as a sink of a tetrahedron \( \sigma \), and \( v \) is the circumcenter of a tetrahedron \( \tau \). Notice that the structure of \( desc(\sigma) \) is a DAG whose node out-degree is at most 2. Consider a tetrahedron \( \tau \) from \( desc(\sigma) \). Without loss of generality, we assume that there is no tetrahedron in \( \pi(\sigma, \tau) \) with a large radius-edge ratio except tetrahedron \( \sigma \) itself. In three dimensions,
the fact that a tetrahedron has a small radius-edge ratio does not guarantee that
the tetrahedron has no small angles. Slivers are the only tetrahedra that have a
small radius-edge ratio but their aspect ratio could be arbitrarily large. Let \( u \) and
\( e_u \) be the circumcenter and the circumradius of \( \tau \). Let \( e_v \) be the circumradius of
the tetrahedron \( \tau \). It is possible that \( e_v \) is almost the same as \( e_u \) even vertex \( v \) is
far away from \( u \). Consequently, \( \frac{e_v - e_u}{||v-u||} \) could not be bounded from below by any
constant. Figure 2 gives an example of a configuration such that \( \frac{e_v - e_u}{||v-u||} \) could be
arbitrarily small. This, together with the fact that \( l_f(s) \) could be as large as
\( l_f(s) + ||v-u|| \) implies that \( D_v = \frac{l_f(s(v))}{e_v} \) could be much larger than \( D_u = \frac{l_f(s(u))}{e_u} \).
Therefore we can not bound the density \( D_v \) using the relation between \( D_u \) and
\( D_v \), even assuming that the density \( D_u \) of vertex \( u \) is bounded by a constant \( D_3 \).

Figure 2 is constructed as follows. Let sliver \( p_0q_0r_0s_0 \) be a successor of the
tetrahedron \( \sigma \), which is not shown in the left figure of Figure 2. Assume that the
circumcenter of the sliver \( p_0q_0r_0s_0 \) is on the different side of the plane \( H_{p_0r_0s_0} \)
passing \( p_0, r_0 \) and \( s_0 \) with sliver \( p_0q_0r_0s_0 \). One of the successor of sliver \( p_0q_0r_0s_0 \)
is constructed by lifting vertex \( q_0 \) to a new position \( q_1 \) such that tetrahedron
\( p_0q_1r_0s_0 \) is a sliver and its circumcenter is on the different side of the plane
\( H_{p_0q_1s_0} \) passing \( p_0, q_1 \) and \( s_0 \) with the tetrahedron \( p_0q_1r_0s_0 \). Then we lift the
node \( r_0 \) to \( r_1 \) to construct a sliver successor \( p_0q_1r_1s_0 \) of tetrahedron \( p_0q_1r_0s_0 \)
whose circumcenter is on the different side of the plane \( H_{p_0q_1r_1} \) with \( p_0q_1r_1s_0 \).
We then construct a sliver successor \( p_0q_1r_1s_1 \) of \( p_0q_1r_1s_0 \) by lifting \( s_0 \) to \( s_1 \). Sliver
successor \( p_1q_1r_1s_1 \) is constructed by lifting \( p_0 \) to a new position \( p_1 \). The above
procedure could be repeated many rounds if \( p_0q_0r_0s_0 \) is carefully configured and
every lifting is carefully chosen. The middle figure in Figure 2 give the sliver
pattern used in constructing this example. It is easy to show that using only the
tetrahedron \( \tau \) also can not bound the density ratio \( D_v = \frac{l_f(s(v))}{e_v} \). Assume that \( p \)

achieved if \( \tau \) is a regular tetrahedron. It implies that the upper bound for \( D_v \)
could be always larger than that for \( D_p \). However, we doubt that these situations
can really happen in practice.

Fig. 2. Left: all sliver descendants; Right: the sink simplex does not help.
5 Conclusion

In this paper, we show that the sink insertion method guarantees to generate a two-dimensional mesh with good grading. On the other hand, we also give an example of three-dimensional local mesh configuration to show that the sink insertion method may fail to generate a mesh with size $O(n)$, where $n$ is the minimum number of the mesh elements with the same radius-edge ratio quality.

As reported by the experimental results (Guoy and Edelsbrunner, private communication), the sink insertion method usually generates meshes whose sizes are not much larger than that by Delaunay refinement method. However, it is interesting to see if we can theoretically prove the good grading guarantee of the sink insertion method or give an example of three-dimensional domain such that a sequence of sink insertions will generate a mesh whose size is larger than $O(n)$. Notice Li [4] recently gave a new refinement-based algorithm that generates well-shaped three-dimensional meshes with size $O(n)$. Li [4] also proposed a variation of the sink insertion method, which inserts a point near the sink and its insertion will not introduce small slivers compared to $\tau$ instead of inserting the sink of a tetrahedron $\tau$ with large radius-edge ratio or sliver. This variation guarantees to generate well-shaped 3D Delaunay meshes. However, it is open whether this variation will have a good grading guarantee. It is interesting to see what is the mesh size relation between two meshes generated by the sink insertion method and this variation proposed in [4].

References

Polynomial Time Algorithms
for Three-Label Point Labeling

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Abstract. In this paper, we present an \(O(n^3 \log n)\) time solution for
the following multi-label map labeling problem: Given a set \(S\) of \(n\)
distinct sites in the plane, place at each site a triple of uniform squares of
maximum possible size such that all the squares are axis-parallel and a
site is on the boundaries of its three labeling squares. We also study the
problem under the discrete model, i.e., a site must be at the corners of
its three label squares. We obtain an optimal \(\Theta(n \log n)\) time algorithm
for the latter problem.

1 Introduction

Map labeling is a popular problem on information visualization in our daily life.
It is an old art in cartography and finds new applications in recent years in GIS,
graphics and graph drawing [AKS98, CMS93, CMS95, DF92, FW91, Imh75,
Jon89], [KR92, KT97, KT98a, PZC98, Wa94, WW95, ZP99]. Among many problems
in map labeling, labeling points is of special interest to many practitioners and
theoreticians. In the paper by Formann and Wagner, any point site can only
be labeled with 4 candidate (axis-parallel) squares each of which has a vertex
anchored at the site and the objective is to maximize the square size. See Figure
1, I.a, for an instance of the problem. (In general we call this kind of model dis-
crete, i.e., each site has \(O(1)\) candidates.) Even this seemingly simplest version
is shown to be NP-complete and moreover; it is NP-hard to approximate within
factor 2 [FW91]. (For details regarding NP-completeness readers are referred to
[GJ79].) In the past several years more generalized models have been proposed.
The basic idea is to allow each site to have an infinite number of possible candi-
date labels (see [DMMZ97, LL97, KSW99, SW99, ZP99]). This model is more
natural than the previous discrete models (like the one in [FW91]) and has been

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coined as the *sliding model* in [KSW99]. On the other hand, designing efficient algorithms for map labeling under the sliding model is a new challenge to map labeling researchers. We briefly review some recent results on labeling points under the sliding model.

Before our review, we briefly define some necessary concepts in approximation algorithms. An approximation algorithm for a (maximization) optimization problem $\Pi$ provides a performance guarantee of $\rho$ if for every instance $I$ of $\Pi$, the solution value returned by the approximation algorithm is at least $1/\rho$ of the optimal value for $I$. For the simplicity of description, we simply say that this is a factor $\rho$ approximation algorithm for $\Pi$.

In [DMMMZ97], Doddi et al. designed several approximation algorithms for labeling points with arbitrarily oriented squares and circles (though the constant factors are impractical: 36.6 and 29.86 respectively). They also presented bicriteria PTAS for the problems. The former bound was improved to 12.95 by Zhu and Qin [ZQ00] and significantly to 5.09 by Doddi et al. most recently [DMM00]. The latter bound on labeling points with circles was improved to 19.35 by Strijk and Wolff [SW99] and recently to 3.6 by Doddi et al. [DMM00]. If any labeling square must be along a fixed direction (e.g., axis-parallel), Zhu and Qin showed that it is possible to have a factor-4 approximation [ZQ00]. In [KSW99] van Kreveld et al. proved that it is NP-hard to decide whether a set of points can all be labeled with axis-parallel unit squares under the sliding model. See Figure 1, I.b, for an example of labeling points with sliding axis-parallel squares. (In fact, in [KSW99], van Kreveld et al. tried to maximize the number of sites labeled instead of the size of the labels.) In [SW99] Strijk and Wolff used similar ideas to prove that the problem of labeling points with maximum size uniform circles is NP-hard. This explains why it is meaningful to study approximation algorithms for these problems.

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**Fig. 1.** Examples for one-label, two-label and three-label point labeling.
As another kind of generalization for map labeling, recently Zhu and Poon studied the problem of labeling point sites with axis-parallel uniform square pairs and circle pairs. The motivation is that in some applications we need two labels for a site \cite{KT98} (like labeling a map for weather reporting or labeling a bilingual city map). They obtained factor-4 and factor-2 algorithms for the two problems respectively, besides presenting a bicriteria approximation scheme \cite{ZP99}. For map labeling with uniform square pairs, Zhu and Qin \cite{ZQ00} first improved the approximation factor of \cite{ZP99} from 4 to 3. Then Qin et al. further improved the factor to 2 \cite{QWXZ00}. More recently, Spriggs proved that the problem is NP-hard and, in fact, NP-hard to approximate within a factor of 1.33 \cite{Sp00}. See Figure 1, II.a and II.b for examples of labeling points with uniform square pairs under the discrete and sliding model respectively.

For map labeling with uniform circle pairs, Qin et al. first improved the 2 approximation factor to 1.96 and proved that problem is NP-hard, in fact, NP-hard to approximate within a constant factor of \( \delta > 1 \) \cite{QWXZ00}. The 1.96 factor was recently improved to 1.686 by Spriggs and Keil \cite{SK00} and then by Wolff et al. to 1.5 \cite{WTX00}. There are still some gaps between the lower and upper bounds for both of the two problems.

In this paper, we study the problem of labeling point sites with uniform square triples. (See Figure 1, III.a, III.b for examples of labeling points with uniform square triples under the discrete and sliding model respectively.) The problem is interesting both in application and theory. In practice, many weather reporting programs on TV need to label a city with three labels: its name, temperature and chance of rainfall. In theory, the problem of labeling point sites with uniform squares is NP-hard under either the discrete model \cite{FW91} or the sliding model \cite{KSW99}. Also, the problem of labeling point sites with uniform square pairs is NP-hard under both the discrete and sliding model \cite{Sp00}. On the other hand, labeling point sites with four squares is trivially polynomial solvable (the solution is decided by the closest pair of the point set, under the \( L_\infty \) metric). The labeling problems for points with circles and circle pairs are both NP-hard \cite{SW99, QWXZ00}. Finally, we remark that it is impossible to label a point with three or more non-overlapping circles.

We now define formally the problems to be studied. We also make some extra definitions related to our algorithms. The MLUST problem (Map Labeling with Uniform Square Triples) is defined as follows:

**Instance:** A set \( S \) of points (sites) \( p_1, p_2, \ldots, p_n \) in the plane.

**Problem:** Does there exist a set of \( n \) triples of axis-parallel squares of maximum size (i.e., length of a side) \( L^* \) each of which are placed at each input site \( p_i \in S \) such that no two squares intersect and no site is contained in any square.

Notice that we can have two versions of the problem: the problem can be under the discrete model in which every site must be at the corners of its three labeling squares and the problem can also be under the sliding model in which every site can be on the boundaries of its three labeling squares. Because of the
nature of the problem even under the sliding model every site must be at the corners of at least two of its labeling squares. We present below a few extra definitions which will be used in later sections.

Given a set $S$ of $n$ sites in the plane, the closest pair of $S$ under $L_\infty$ metric, $D_\infty(S)$, is defined as the minimum $L_\infty$-distance between any two points in $S$. Clearly $D_\infty(S)$ can be computed in $O(n \log n)$ time with standard algorithm in computational geometry [PS85].

2 MLUST under the Discrete Model

In this section we present the details of a polynomial time algorithm for the MLUST problem under the discrete model. Because of the nature of the problem, the metric discussed in this paper is $L_\infty$ unless otherwise specified. Let $D_\infty(S)$ be the closest pair of $S$ under the $L_\infty$-metric. Let $l^*$ denote the size of each square in the optimal solution of the discrete MLUST problem. The following lemma is easy to prove.

Lemma 1. $D_\infty(S)/2 \leq l^* \leq D_\infty(S)$.

In the following we show how to decide whether a set of points $S$ can be labeled with square triples with edge length $l$, $D_\infty(S)/2 \leq l \leq D_\infty(S)$. For any two points $p_i, p_j \in S$, let $d_\infty(p_i, p_j)$ denote the $L_\infty$-distance between them. Let $C_i$ denote the $L_\infty$-circle centered at point $p_i \in S$ with radius $l$. Clearly, the circle $C_i$ contains no other point from the input set $S$ except its own center. Note that each circle $C_i$ can be partitioned into four $L_\infty$-circles with radius $l/2$, which are geometrically squares with edge length $l$. We loosely call them sub-squares of $C_i$. Basically, to label $p_i$ we need to select three out of four sub-squares in each $C_i$ so that no two sub-squares intersect each other.

We now present our algorithm. First, we compute the following multiple intersection graph $G_M(S, l)$. $C_i$ ($1 \leq i \leq n$) are the vertices for $G_M(S, l)$. There is an edge between $C_i, C_j$ if they intersect and only one pair of sub-squares of them overlap. If at least two (and at most three) pairs sub-squares of $C_i, C_j$ overlap, then we draw a multiple of two edges between $C_i, C_j$. (Recall that two squares overlap if they have a common interior point.)

Lemma 2. There is a valid labeling for $S$ with square triples of size $l$ if and only if every connected component of $G_M(S, l)$ has at most one cycle.

Proof: We refer to Figure 2. By the definition of the problem, if we have a cycle in $G_M(S, l)$ which contains no multiple edge then when we label $p_i$ the labeling will generate ‘pressure’ to either $C_j$ or $C_k$. In other words, one of the sub-squares of either $C_j$ or $C_k$ will be ‘destroyed’ and cannot be used as legal label for either $p_j$ or $p_k$ anymore. This holds for all the nodes involved in that cycle. (In Figure 2, Case 1, the labeling of $p_i$ generates ‘pressure’ on $C_j$.) If $C_i, C_j$ form a cycle with two multiple edges, similar claim holds, except that we need to label $p_i$ ($p_j$) with a sub-square which only generates one ‘pressure’ to $C_j$ ($C_i$). (See Figure 2, Case 2.) Now we continue with our proof.
(Necessity.) Assume that there is a valid labeling for $S$ with square triples of size $l$. Then for each $C_i$, it receives at most one 'pressure', i.e., at most one of its four sub-squares cannot be used as legal label for $p_i$. Clearly, that implies that every connected component of $G_M(S, l)$ contains at most one cycle.

(Sufficiency.) It is easy to see that a leaf dangling at a cycle has no influence over the labeling of the centers of those nodes involved in a cycle. So we assume that there is no leaf node in any connected component of $G_M(S, l)$. Now if each connected component of $G_M(S, l)$ is a cycle then obviously we can label the centers of all those nodes $C_i$ involved in that cycle. □

Clearly the graph $G_M(S, l)$ has a vertex degree of at most 12. So the graph is of linear size and checking whether the graph contains more than one cycle can be done in linear time with standard graph algorithm. To obtain a polynomial time solution for MLUST under the discrete model all we need to do is to prove that there are only a polynomial number of candidates for $l$. Among them, the largest will give us the size of the optimal solution. Let the coordinates of $p_i$ be $(x(p_i), y(p_i))$ and let $d_{\min}(p_i, p_j) = \min\{|x(p_i) - x(p_j)|, |y(p_i) - y(p_j)|\}$. (Note that $d_{\infty}(p_i, p_j) = \max\{|x(p_i) - x(p_j)|, |y(p_i) - y(p_j)|\}$.) We have the following lemma.

**Lemma 3.** The size of the optimal solution for MLUST under the discrete model is equal to either $d_{\infty}(p_i, p_j)$ (in which case it is exactly $D_{\infty}(S)$), or $d_{\infty}(p_i, p_j)/2$ or $d_{\min}(p_i, p_j)$; provided that its value is bounded by $D_{\infty}(S)/2$ and $D_{\infty}(S)$.

Lemmas 1, 2 and 3 naturally give us the following algorithm. For each site $p_i$, we look at the axis-parallel square $C$ centered at $p_i$ with edge length $4D_{\infty}(S)$. Clearly any point $p_k$ out of this square would be at a distance longer than $2D_{\infty}(S)$, which implies $d_{\infty}(p_i, p_k)/2$ cannot be the optimal solution value for the problem. As any two sites in $C$ must be at least $D_{\infty}(S)$ distance away, we only need to consider a constant number (24) points in $S$ which are the closest to $p_i$. (Overall, for all $p_i$ this can be computed in $O(n \log n)$ time using standard techniques [DLSS93].) For each such point $p_j$, we simply measure $d_{\infty}(p_i, p_j)$ and $d_{\min}(p_i, p_j)$. If $d_{\infty}(p_i, p_j)/2$ or $d_{\min}(p_i, p_j)$ is out of the range $[D_{\infty}(S)/2, D_{\infty}(S)]$
then throw it away as a valid candidate. Eventually we have at most \(2 \times 24n + 1 = O(n)\) number of candidates. We sort them into a list \(l_1, \ldots, l_{O(n)}\) and then we run a binary search over this list to decide the maximum value \(l^*\) such that a valid labeling for \(S\) with such a value, exists. As the decision step, following Lemma 2, takes \(O(n)\) time, the whole algorithm takes \(O(n \log n)\) time. It is easy to show that \(\Omega(n \log n)\) is the lower bound for the MLUST problem under either the discrete or the sliding model, by a reduction from the element uniqueness problem. Summarizing the above results, we have the following theorem.

**Theorem 1.** For any given set of \(n\) points in the plane, the above algorithm, which runs in \(\Theta(n \log n)\) time, produces an optimal solution for the MLUST problem under the discrete model.

### 3 MLUST under the Sliding Model

In this section, we shall proceed with the more interesting problem of labeling a set of points with sliding square triples. Because of the nature of the problem, not all of the three labels for a site \(p_i\) could slide; in fact, only one of them could. We hence call the two discrete sub-squares in the label of \(p_i\) base sub-squares or base labels. Let \(L^*\) be the optimal solution of the problem MLUST under the sliding model. We have a lemma similar to Lemma 1.

**Lemma 4.** \(D_\infty(S)/2 \leq L^* \leq D_\infty(S)\).

Our general idea is the same as that for the discrete case. We first try to design a decision procedure which can decide for any \(l \in [D_\infty(S)/2, D_\infty(S)]\) whether a valid labeling of \(S\) with sliding square triples of size \(l\) exists. We follow the same procedure in the previous section to build a multi-graph \(G_M(S, l)\). (The nodes of \(G_M(S, l)\) are the set of \(C_i\) whose centers are \(p_i\), for all \(p_i \in S\).) However, because of the difference between the two models we cannot immediately have a lemma similar to Lemma 2. The reason is that the sliding of some label for \(p_i\) might simply terminate any ‘pressure’ its neighbor carries over to it. In Figure 3 (a), if we label \(p_i\) using the shaded sliding label then the ‘pressure’ from \(p_k\) vanishes and

![Fig. 3. A pressure-releasing operation.](image-url)
in Figure 3 (b), if we label $p_i$ using the shaded sliding label then the ‘pressure’ from $p_j$ vanishes. We call $p_j, p_i, p_k$ a critical triple if $C_j$ intersects $C_i$ and $C_i$ intersects $C_k$. A pressure-releasing operation on a critical triple $p_j, p_i, p_k$ is that we label $p_i$ with sliding labels such that the labels generate minimum ‘pressure’ on either $p_j$ or $p_k$, i.e., the number of either $p_j$ or $p_k$’s sub-squares destroyed by the labels of $p_i$ is minimized and furthermore, the area of both $p_j$ and $p_k$’s sub-squares destroyed by the labels of $p_i$ is also minimized. (The first condition implies that a pressure-releasing operation destroys either 0 or 1 sub-square of $p_j$ or $p_k$. The second condition implies that if a pressure-releasing operation has to destroy a sub-square of $p_j$ or $p_k$ then it will destroy the minimum area of it. Finally, it is clear that we can perform $O(1)$ number of pressure-releasing operations on any given critical triple.) In this case we call $C_i$ a cycle-breaker in $G_M(S, l)$ and clearly a cycle-breaker will terminate some pressure along some cycle in $G_M(S, l)$ if we perform a pressure-releasing operation on its center. Therefore, we have the following revised version of Lemma 2, whose proof is straightforward.

**Lemma 5.** There is a valid labeling for $S$ with sliding square triples of size $l$ if and only if every connected component of $G_M(S, l)$ has at most one cycle after a set of pressure-releasing operations are enumerated.

Notice that different from the discrete MLUST problem, this lemma does not give us a static algorithm as does Lemma 2. In fact, at the first glance, it seems that the above lemma gives us an exponential solution. However, we will make use of the specific property of $G_M(S, l)$ to obtain a polynomial time solution.

What we do is as follows. As the vertex degree of any node in $G_M(S, l)$ is at most a constant (12), we can fix any node $C_i$ and identify all possible states of it. That state is determined by the position of the two base sub-squares as well as along which path adjacent to $C_i$ the ‘pressure’ will be generated. Clearly we have $O(1)$ number of states for $C_i$.

What we do next is to fix a state of $C_i$ and traverse the graph by following those paths carrying ‘pressures’ generated so far. Suppose that after visiting $C_j$ and successfully labeling $p_j$ (i.e., the current state of $p_j$ is safe), we reach at a vertex $C_k$. If we can find a valid labeling of $p_k$ taking into consideration all the pressures generated on $p_k$ so far, then we set the state of $p_k$ as safe and we continue our traversal. If we reach a dead state at $p_k$, i.e., no valid labeling of $p_k$ exists (in other words, at least two of $p_k$’s sub-squares are destroyed), then we backtrack to $C_j$ and traverse the edge $(C_j, C_k)$ by starting at a different safe state of $p_j$, if there exists one. If we backtrack to $p_i$ and try out all its safe states and still cannot find a valid labeling for all the sites in $S$, then a valid labeling of $S$ with square triples of size $l$ does not exist.

At the first sight, this procedure seems to take exponential time. However, the following lemma guarantees a polynomial time solution for fixed $l$. (It is obvious that if we can label all sites corresponding to nodes in $G_M(S, l)$ with square triples of size $l$ then there must exist at least one node in $G_M(S, l)$ which admits a pressure-releasing operation.)
Lemma 6. In the above procedure, if at each cycle-breaker $C_j$ we minimize the out-going pressure along $(C_j, C_k)$ while withholding the incoming ‘pressure’ then for each state of $C_j$ there is a unique state for the next cycle-breaker $C_k$ in the same cycle.

The above lemma basically shows that the procedure in the previous paragraph runs in linear time if we start at a $C_i$ which admits a pressure-releasing operation. (For each cycle-breaker $C_j$ we have $O(1)$ state, when traversing the graph from $C_j$ to the next cycle-breaker $C_k$ we can only reach exactly one state of $C_k$. This procedure finishes either after we try all states of $C_i$ without finding a valid labeling for $S$ or we terminate with a valid labeling of size $l$ for set $S$.) However, as in the optimal labeling not all $C_i$’s admit a pressure-releasing operation we need to try the above procedure $O(n)$ times — starting at every possible node in the graph. Therefore, for a fixed $l$ deciding whether we can label $S$ with sliding square triples of size $l$ can be done in $O(n \times n) = O(n^2)$ time.

To solve MLUST under the sliding model, we must also make sure that there are only a polynomial number of candidates for $L^*$. This is guaranteed with the following lemma.

Lemma 7. The size of the optimal solution for MLUST under the sliding model $L^*$ is equal to either the optimal solution for MLUST under the discrete model or $d_\infty(p_i, p_j)/K$ for some $p_i, p_j \in S$ and some $K$ such that $1 \leq K \leq n - 2$, provided that its value is bounded by $D_\infty(S)/2$ and $D_\infty(S)$.

Proof. It is only necessary to discuss the situation when $L^*$ is not equal to the optimal solution for MLUST under the discrete model. In this situation, the reason that the optimal solution value $L^*$ for MLUST under the sliding model cannot be increased is that there exists a series of $K$ labels $p'_1, ..., p'_K$ touching each other and the sum of their sizes is exactly the distance between some sites $p_i, p_j$, i.e., each $p'_k$ contributes some distance to fill the $L_\infty$-gap between $p_i, p_j$ (Figure 4). We call $p_i, p_j$ an extreme pair. (Note that in Figure 5 the base labels for all the sites are not shown.) What remains to show is that each $p'_k$ contributes exactly a distance of $L^*$ to fill the $L_\infty$-gap between $p_i, p_j$, which in turn implies that we do not need to consider any $K$ which is larger than $n - 2$. Assume to the contrary that this is not the case, i.e., at least one of the sites, say $p'_k$, would contribute $2L^*$ to fill the $L_\infty$-gap between $p_i, p_j$. However, this implies that either $p_i, p'_k$ or $p'_k, p_j$ would be an extreme pair. □

Similar to the previous section, Lemmas 4, 5, 6 and 7 naturally give us the following algorithm. For each pair of sites $p_i, p_j$, we simply measure $d_\infty(p_i, p_j)$. If any of $d_\infty(p_i, p_j)/K (1 \leq K \leq n - 2)$ is out of the range $[D_\infty(S)/2, D_\infty(S)]$ then we throw it away as a valid candidate. With $i, j$ fixed, we have at most $O(n)$ number of candidates. In total we have $O(n^3)$ number of candidates for $L^*$. (We also need to test all the $O(n)$ candidates for the discrete problem.) We sort them into a list $L_1, ..., L_{O(n^3)}$ in $O(n^2 \log n)$ time and then we run a binary search over this list to decide the maximum value $L^*$ such that a valid labeling for $S$ with such a value, exists. As the decision step, following Lemma 6, takes
Fig. 4. Illustration for the proof of Lemma 7.

$O(n^2)$ time, the whole algorithm takes $O(n^3 \log n + n^2 \times \log n^3) = O(n^3 \log n)$ time. Summarizing the above results, we have the following theorem.

**Theorem 2.** For any given set of $n$ points in the plane, there is an $O(n^3 \log n)$ time solution for the MLUST problem under the sliding model.

4 Concluding Remarks

In this paper, we present polynomial time algorithms for the problem of labeling point sites with uniform square triples, under either the discrete or sliding models. This is significantly different from the problem of labeling point sites with uniform square pairs, which is NP-hard under both the discrete and sliding models. An immediate question is whether we can reduce the gap between the $\Omega(n \log n)$ lower bound and the $O(n^3 \log n)$ upper bound for the general problem (under the sliding model).

References


Approximation Algorithms for the Watchman Route and Zookeeper’s Problems

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1 Preliminary

Given a simple polygon $P$ with $n$ vertices and a starting point $s$ on its boundary, the watchman route problem asks for a shortest route in $P$ through $s$ such that each point in the interior of the polygon can be seen from at least one point along the route. It is known that the watchman route problem can be reduced in $O(n \log n)$ time to that of computing the shortest route which visits a set of line segments in polygon $P$. In this paper, we present a simple approximation algorithm for computing the shortest route visiting that set of line segments. Our algorithm runs in $O(n)$ time and produces a watchman route of at most 2 times the length of the shortest watchman route. The best known algorithm for computing the shortest watchman through $s$ takes $O(n^4)$ time. Our scheme is also employed to give a $\sqrt{2}$-approximation solution to the zookeeper’s problem, which is a variant of the watchman route problem.

The known watchman route and zookeeper’s algorithms primarily make use of the reflection principle. If $a$ and $b$ are two points on the same side of a line $L$, then the shortest path visiting $a$, $L$ and $b$ in that order, denoted by $S(a, L, b)$, follows the reflection principle, i.e., the incoming angle of path $S(a, L, b)$ with $L$ is equal to the outgoing angle of $S(a, L, b)$ with $L$. To actually compute the path $S(a, L, b)$, we first reflect $b$ across $L$ to get its image $b'$, then draw the straight segment $ab'$ from $a$ to $b'$, and finally fold back the portion of $ab'$ lying in the other side of $L$ to obtain the path $S(a, L, b)$. See Fig. 1a for an example. More generally, to find a shortest path from $a$ to $b$ that visits a segment $l$, denoted by $S(a, l, b)$, we first compute the shortest path that passes between the endpoints of segment $l$ and connects $a$ to the reflection $b'$ and then fold back the portion of the path in the other side of $l$ (Fig. 1b).

Let $L(a)$ denote the point of $L$ nearest to $a$ (Fig. 1). The path consisting of the segments $aL(a)$ and $L(a)b'$, denoted by $S'(a, L, b)$, is a good approximation of $S(a, L, b)$. Note that $a$, $L(a)$ and $b'$ form an obtuse-angled triangle. Since the sum of lengths of two shorter edges of a right-angled triangle is the maximum when two acute angles are $\pi/4$, we have the following result.

**Lemma 1** For an obtuse-angled triangle, the sum of lengths of two shorter edges is smaller than or equal to $\sqrt{2}$ times the length of the longest edge.
Analogously, let \( l(a) \) denote the point of the segment \( l \) nearest to \( a \) (Fig. 1b). Then the path consisting of two segments \( al(a) \) and \( lb(a)b \), denoted by \( S'(a, l, b) \), is also a \( \sqrt{2} \)-approximation of the path \( S(a, l, b) \).

2 Approximating the Shortest Watchman Route

Let \( P \) be a polygon with a point \( s \) on its boundary. A vertex is reflex if its internal angle is greater than \( \pi \). Polygon \( P \) can be partitioned into two pieces by a "cut" that starts at a reflex vertex \( v \) and extends either edge incident to \( v \) until it first hits the boundary. We say a cut is a visibility cut if it produces a convex angle \( (< \pi) \) at \( v \) in the piece of \( P \) containing \( s \). For a cut \( C \), the piece of \( P \) containing \( s \), denoted by \( P(C) \), is called the essential piece of \( C \). We say cut \( C_j \) dominates cut \( C_i \) if \( P(C_j) \) contains \( P(C_i) \). A cut is called an essential cut if it is not dominated by any other cuts. Clearly, a route is a watchman route if and only if it visits all essential cuts. (All essential cuts can be computed in \( O(n \log n) \) time.)

Let \( C_1, C_2, \ldots, C_m \) the sequence of essential cuts indexed in the clockwise order of their left endpoints, starting at \( s \). Also, let \( s = s_0 = s_{m+1} \). First, we find the point of \( C_1 \) that is closest to \( s_0 \) in \( P(C_1) \). Let \( s_1 \) denote the point found on \( C_1 \), which is called as the image of \( s_0 \) on \( C_1 \). Similarly, we can find the images of \( s_0 \) on \( C_2, C_3 \) and so on. In the case where \( s_k \) is not contained in \( P(C_{k+1}) \), the image \( s_{k+1} \) of \( s_k \) on \( C_{k+1} \) is undefined, and the image \( s_{k+2} \) of \( s_0 \) on \( C_{k+2} \) is considered instead. The computation of images of \( s_0 \) is terminated when a new image \( s_{i+1} \) does not dominate all the cuts before it (i.e., \( C_1, \ldots, C_i \)). Next, we take \( s_i \) as a new starting point and compute the images of \( s_i \) on the cuts after it. This procedure is repeatedly performed, until the image on \( C_m \) is computed (or considered). In particular, we call the images, which are considered as the starting points, the critical images. See Fig. 2 for an example where images \( s_0 (= s_6), s_2, s_4 \) and \( s_5 \) are critical. (Image \( s_{m+1} \) is not shown in our figure.)

Let \( R' \) denote the route which is a concatenation of the shortest paths between every pair of adjacent critical images \( s_i \) and \( s_j \) \( (0 \leq i < j \leq m + 1) \). Clearly, \( R' \) is a watchman route, which gives our approximation solution. Let \( R \) denote the shortest watchman route through \( s \). An example of routes \( R' \) and \( R \) is shown in Fig. 2. In the following, we denote by \( |xy| \) the length of a line segment \( xy \), and \( |Z| \) the length of a route \( Z \).
A watchman route $R$ is said to make a reflection contact with cut $C$ if $R$ comes into $C$ at some point and then reflects on $C$ and goes away from that point.

**Lemma 2** If route $R$ reflects on cut $C_i$ and the critical image $s_i$ is defined, then $s_i$ is to the left of the reflection point of $R$ on $C_i$ (as viewed from $s_0$).

**Proof.** Omitted in this extended abstract (see also Fig. 2 for an example). $\square$

**Theorem 1** For any instance of the watchman route problem, $|R'| \leq 2|R|$. Moreover, the watchman route $R'$ can be found in $O(n)$ time, provided that the set of essential cuts is given.

**Proof.** Assume first that route $R'$ does not touch the boundary of the given polygon $P$ (except at $s$). Let $s_i$ and $s_j$ ($1 \leq i < j \leq m$) denote the first two critical images (excluding $s_0$). If $s_j$ is contained in $P(C_i)$, it then follows from Lemma 2 that two routes $R$ and $R'$ intersect before $s_j$. Let $t$ be the intersection point immediately before $s_j$ (along the route $R'$), and $R_{0,t}$, $R'_{0,t}$ the parts of routes $R$, $R'$ from $s_0$ to $t$, respectively. Let $t'$ denote the point obtained by reflecting $t$ across $C_i$. Since the chain of segments $s_0s_i$ and $s_it'$ gives the unfolded route $R'_{0,t}$, we have that $|s_0t'| \leq |R_{0,t}|$. On the other hand, since three points $s_0$, $s_i$ and $t'$ form an obtuse-angled triangle, it follows from Lemma 1 that $|R'_{0,t}| \leq \sqrt{2}|s_0t'|$, and thus $|R'_{0,t}| \leq \sqrt{2}|R_{0,t}|$.

Consider the situation where $s_j$ is not contained in $P(C_i)$. Let $t$ be the first intersection of route $R'$ with $R$, immediately after $s_j$. (See also Fig. 2 for an example where $s_i = s_2$ and $s_j = s_4$.) Since the next critical image is contained in $P(C_j)$ (otherwise, $s_j$ were not critical), point $t$ is contained in $P(C_j)$, too. Let $t'$ denote the point obtained by reflecting $t$ across $C_j$. Since $s_i$, $s_j$ and $t'$ form an obtuse-angled triangle, $|R'_{i,t}| \leq \sqrt{2}|s_it'|$. Let $R'_{0,t'}$ denote a new route that consists of $s_0s_i$ and $s_it'$. Then we have that $|R'_{0,t'}| \leq \sqrt{2}|s_0t''|$ and $|s_0t''| \leq \sqrt{2}|R_{0,t}|$, where $t''$ is the point obtained by reflecting $t'$ across $C_i$ if $t'$ is contained in $P(C_i)$ (see also Fig. 2), or $t'' = t'$ otherwise. Since $|R'_{0,t'}| = |s_0s_i| + |s_it'| \geq |s_0s_i| + \frac{\sqrt{2}}{2}|R'_{i,t}| > \frac{\sqrt{2}}{2}(|s_0s_i| + |R'_{i,t}|) = \frac{\sqrt{2}}{2}|R_{0,t}|$, we have that $|R'_{0,t'}| < 2|R_{0,t}|$. 

**Fig. 2.** Critical images and routes $R'$, $R$. 
The rest parts of $R$ and $R'$ can inductively be considered by taking $t$ as a new starting point. Note that if the rest part of $R'$, say, $R'_{t,m+1}$, contains only one critical image (excluding $s_{m+1}$), then $|R'_{t,m+1}| \leq \sqrt{2}|R_{t,m+1}|$.

Let us now remove the assumption that $R'$ does not touch the boundary of polygon $P$. Consider first the case where $s_i$ is the left endpoint of $C_i$. If the incoming angle of $R'$ with $C_i$ is larger than the outgoing angle (Fig. 3a), the first part of the theorem is still true. Otherwise, it follows from Lemma 2 that $s_i$ should be visited by the route $R$ (Fig. 3b). Then, the first part of the theorem can inductively be proved for two parts of $R$ and $R'$ separated by $s_i$ (i.e., one starts at $s_0$ and ends at $s_i$, and the other starts at $s_i$ and ends at $s_{m+1}$). The other case having to be dealt with is that the shortest path between two critical images, say, $s_i$ and $s_{i+1}$, wraps around some reflex vertices of $P$ (i.e., the path makes left turns at these vertices when one follows it in the clockwise direction). See Fig. 3c. Let $r$ and $r'$ denote the last and penultimate reflex vertices of $P$ touched by the shortest path between $s_i$ and $s_{i+1}$. Note that $r'$ is the first point of the shortest path between $s_i$ and $s_{i+1}$, which is visible to some portion of $C_{i+1}$. It follows from Lemma 2 that $R$ has to touch $r'$ and/or $r$. If $r$ is visited by $R$, the first part of the theorem can inductively be proved for two parts of $R'$ and $R$, separated by the vertex $r$. In the situation where only $r'$ is visited by $R$, the similar treatment can be done for the parts of $R$ and $R'$ from $s_0$ to $r'$. Let $R_{r',m+1}$ and $R'_{r',m+1}$ denote the parts of $R$ and $R'$ from $r'$ to $s_{m+1}$, respectively. Since $r$ is a reflex vertex and $s_{i+1}$ is invisible to $r'$, we extend the segment $rs_{i+1}$ until the foot of a perpendicular to the extension and through $r'$ is reached (Fig. 3d). Let $R'_{r',m+1}$ denote the enlarged route. It is easy to see that $|R'_{r',m+1}| < |R''_{r',m+1}| \leq 2|R'_{r',m+1}|$. It completes the proof of the first part of the theorem.

Finally, the watchman route $R'$ can be computed in $O(n)$. Due to space limit, we omit the detail. ✷

Fig. 3. Illustration for the proof of Theorem 1.
3 Approximating the Shortest Zookeeper’s Route

The zookeeper’s problem is defined as follows: Given a starting point $s$ and a simple polygon (zoo) $P$ with a set $\mathcal{P}$ of disjoint convex polygons (cages) inside it, each sharing one edge with $P$, we want to find the shortest route $R$ through $s$ that visits (without entering) at least one point of each cage in $P$. The best algorithm for computing the shortest zookeeper’s route takes $O(n \log^2 n)$ time [1]. Jonsson also gave a 6-approximation solution to the zookeeper’s problem [2].

Let $P_1, \ldots, P_m$ denote the cages indexed in a clockwise scan of the boundary of zoo $P$, starting at $s$. Also we find a point (the image) on the boundary of $P_1$ that is closest to $s_0$ in the interior of zoo $P$ (i.e. $P-P$), a point $s_2$ on the boundary of $P_2$ that is closest to $s_1$, and so on. Putting the shortest paths between $s_i$ and $s_{i+1}$ for all $i$, $0 \leq i \leq m$, gives a zookeeper’s route $R'$.

**Theorem 2** For any instance of the zookeeper’s problem, $|R'| \leq \sqrt{2}|R|$. Moreover, the zookeeper’s route $R'$ can be found in $O(n)$ time.

**Proof.** Our proof is done by constructing a route $R''$ such that $|R'| \leq \sqrt{2}|R''|$ and $|R''| \leq |R|$. Let $e_i$ be the edge of cage $P_i$ containing the image $s_i$. Project all vertices of $P_i$ onto the line containing $e_i$, and then extend edge $e_i$ in the interior of zoo $P$ rightward (as viewed from $s$) until the extreme of the projected vertices is reached (or as longer as possible otherwise). See Fig. 4a for an example. Let $E$ denote the set of the line segments, which are obtained by extending all image-defined edges, and let $R''$ be the shortest route computed by the unfolded method, using the line segments in $E$ as mirrors. (Route $R''$ may not be a zookeeper’s route.) Since $R'$ can also be considered as the route reflecting on all elements of $E$, it follows the proof of Theorem 1 that $|R'| \leq \sqrt{2}|R''|$. 

![Fig. 4. Illustration for the proof of Theorem 2.](image)

Assume that $a$ and $b$ are two points on the same side of the edge $e$ of a convex polygon $Q$ with $a.x < b.x$ and that the foot of a perpendicular through $a$ is contained in $e$ (Fig. 4b). Let $e'$ denote the line segment obtained by extending edge $e$ rightward until the extreme of the projected vertices of $Q$ is reached, and $S(a, e', b)$ the shortest path visiting $a$, $e'$ and $b$ in that order. Let $S(a, Q, b)$ denote the shortest path that visits $a$, some point on the boundary of $Q$, and $b$ in that order. Let $c$ denote the first intersection of $S(a, Q, b)$ with $e'$. Clearly, $|S(a, e', b)| \leq |ac| + |cb| \leq |S(a, Q, b)|$ (Fig. 4b). It implies that $|R''| \leq |R|$. Again, the time taken to compute the route $R'$ is $O(n)$. □
References

Abstract. A data structure called PC-tree is introduced as a generalization of PQ-trees. PC-trees were originally introduced in a planarity test of Shih and Hsu [7] where they represent partial embeddings of planar graphs. PQ-trees were invented by Booth and Lueker to test the consecutive ones property in matrices. The original implementation of the PQ-tree algorithms by Booth and Lueker using nine templates in each bottom-up iteration is rather complicated. Also the complexity analysis is rather intricate. We give a very simple PC-tree algorithm with the following advantages: (1) it does not use any template; (2) it does all necessary operations at each iteration in one batch and does not involve the cumbersome bottom-up operation. PC-trees can be used naturally to test the circular ones property in matrices. And the induced PQ-tree algorithm can considerably simplify Booth and Lueker’s modification of Lempel, Even and Cederbaum’s planarity test.

1. Introduction

A data structure called PC-tree is considered here as a generalization of PQ-trees. PC-trees were originally introduced to represent partial embeddings of planar graphs in Shih and Hsu [7]. PQ-trees were used to test the consecutive ones property in matrices [1]. However, the implementation of PQ-tree algorithms by Booth and Lueker [1] (hereafter, referred to as the B&L algorithm) is rather complicated. Also the complexity analysis is rather intricate. We shall present a very simple PC-tree algorithm without using any template. Furthermore, we shall illustrate how to test the circular ones property in matrices using PC-trees.

PQ-trees were invented for the more general purpose of representing all permutations of a set $U$ that are consistent with constraints of consecutiveness given by a collection $C$ of subsets of $U$ with the convention that the element of each subset $S$ in $C$ must occur consecutively in the permutation.

The literature on problems related to PQ-trees is quite extensive. PQ-trees can be applied to test the consecutive ones property of $(0,1)$-matrices [1,3], to recognize interval graphs [1,2] and to recognize planar graphs efficiently [1,6]. Korte and Möhring [5] considered a modified PQ-tree and a simpler incremental update of the tree for the purpose of recognizing interval graphs. Klein and Reif [4] constructed efficient parallel algorithms for manipulating PQ-trees. On the other hand, PC-trees were initially used in Shih and Hsu [7] (S&H) to represent partial embeddings of planar graphs. Note that this approach is entirely different from Lempel, Even and...
Cederbaum’s (LEC) approach [6] of using PQ-trees to test the consecutive ones property of all nodes adjacent to the incoming node in their vertex addition algorithm. In S&H, a P-node is a regular node of the graph, a C-node represents a biconnected component and nodes adjacent to the node in consideration can be scattered anywhere around the PC-tree. However, in LEC, the leaves of the PQ-tree must be those nodes adjacent to the incoming node.

In this paper we shall focus on the application of PQ-trees to (0,1)-matrices. A (0,1)-matrix \(M\) has the consecutive ones property (COP) for the rows iff its columns can be permuted so that the ones in each row are consecutive. \(M\) is said to have the circular ones property (CROP) for the rows if either the ones or the zeros of each row are consecutive. When the zeros of a row are consecutive the ones will wrap around the first and the last column.

A PQ-tree is a rooted tree \(T\) with two types of internal nodes: \(P\) and \(Q\), which will be represented by circles and rectangles, respectively. The leaves of \(T\) correspond 1-1 with the columns of \(M\).

\[
\begin{array}{ccccccc}
1 & 2 & 3 & 4 & 5 & 6 \\
1 & 1 & 0 & 0 & 0 & 0 \\
0 & 1 & 1 & 0 & 0 & 0 \\
1 & 1 & 1 & 1 & 0 & 0 \\
0 & 0 & 0 & 1 & 1 & 1 \\
0 & 0 & 0 & 0 & 1 & 1 \\
\end{array}
\]

\[Q\]
\[P\]

\[1 \quad 2 \quad 3 \quad 4 \quad 5 \quad 6\]

\[Q\]
\[P\]

\[1 \quad 2 \quad 3 \quad 4 \quad 5 \quad 6\]

\[\text{Fig. 1. A PQ-tree and the consecutive ones property}\]

We shall follow the notations used in [1]. The frontier of a PQ-tree \(T\), denoted by \(\text{FRONTIER}(T)\), is the permutation of the columns obtained by ordering the leaves of \(T\) from left to right. Such a permutation is called a consistent permutation. The B&L algorithm considers each row as a constraint and adds them one at a time. Each time a new row is added, the algorithm tries to modify the current PQ-tree to satisfy the consecutiveness constraint of columns of the newly added row. To guarantee unique PQ-tree representations we need to restrict ourselves to proper PQ-trees defined in [1]: every \(P\)-node has at least two children; every \(Q\)-node has at least three children.

Two PQ-trees are equivalent iff one can be transformed into the other by applying zero or more equivalent transformations. The equivalence of two trees is written \(T \sim T'\). There are two types of equivalent transformations:

1. Arbitrarily permute the children of a \(P\)-node,
2. Reverse the order of the children of a \(Q\)-node.

Denote the set of all consistent permutations (or frontiers) of a PQ-tree \(T\) by \(\text{CONSISTENT}(T) = \{\text{FRONTIER}(T') \mid T' \sim T\}\). Given a PQ-tree \(T\) and a row \(u\) of \(M\), define a new tree called the \(u\)-reduction of \(T\) to be one whose consistent permutations are exactly the original permutations in which the leaves in \(u\) occur consecutively. This new tree is denoted by \(\text{REDUCE}(T,u)\). Booth and Lueker gave a procedure \(\text{REDUCE}\) to obtain the \(u\)-reduction. The procedure applies a sequence of templates to the nodes of \(T\). Each template has a pattern and a replacement. If a node matches the template’s pattern, the pattern is replaced within the tree by the
template’s replacement. This is a bottom-up strategy that examines the tree node-by-node obeying the child-before-parent discipline.

A node is said to be full if all of its descendants are in \( u \); it is empty if none of its descendants are in \( u \); if some but not all of the descendants are in \( u \), it is said to be partial. Nodes are said to be pertinent if they are either full or partial. The pertinent subtree of \( T \) with respect to \( u \), denoted by \( \text{PERTINENT}(T, u) \), is the subtree of minimum height whose frontier contains all columns in \( u \). The root of the pertinent subtree is denoted by \( \text{ROOT}(T, u) \), which is usually not the root of the entire tree. There are nine templates used in [1].

In order to achieve optimal efficiency, the algorithm takes several precautions in scanning the pertinent subtree. For example, the maintenance of parent pointers may cause a problem. It is possible that almost all nodes in the tree may receive a new parent even though the row \( u \) has only two elements. Therefore, parent pointers are only kept for children of \( P \)-nodes and for endmost children of \( Q \)-nodes. Interior children of a \( Q \)-node do not keep parent pointers. Rather, they borrow them from their endmost siblings. These measures unavoidably complicate the implementation of the B&L algorithm.

A \( PC \)-tree is a rooted tree \( T \) with two types of internal nodes: \( P \) and \( C \), which will be represented by circles and double circles, respectively. The leaves of \( T \) correspond 1-1 with the columns of \( M \). Figure 2 gives an example of a matrix satisfying the \( CROP \) (but not the \( COP \)) and its corresponding \( PC \)-tree. In obtaining consistent permutations, the children of a \( P \)-node can be permuted arbitrarily whereas the children of a \( C \)-node observe a circular list, whose order can only be changed from clockwise to counter-clockwise. The \( \text{FRONTIER}(T) \) of a \( PC \)-tree consists of those consistent circular permutations.

For each \( PQ \)-tree, we can obtain its corresponding \( PC \)-tree by replacing the \( Q \)-nodes with \( C \)-nodes. Note that flipping a \( Q \)-node in a \( PQ \)-tree is the same as changing the circular children list of its corresponding \( C \)-node from clockwise to counter-clockwise order. There are a few differences between \( PC \)-trees and \( PQ \)-trees: (1) When a \( PQ \)-tree is transformed into a \( PC \)-tree, and its root is a \( Q \)-node, then its two endmost children must be preserved in order to maintain the \( COP \) in any consistent circular permutation of the transformed \( PC \)-tree. Aside from this restriction, all other \( PC \)-tree operations will yield equivalent \( PQ \)-tree operations. (2) Since it is the circular order of the leaves that needs to be preserved (for example, in Figure 2, we can rotate the tree so that \( \text{FRONTIER} \) becomes 2,3,4,5,6,1), one might as well fix the first column to be 1 in a \( PC \)-tree in considering any consistent circular permutation. (3)
The root of a PC-tree in general is not essential except to maintain the child-parent relationships.

With the additional freedom of rotation, it is natural for PC-trees to test the COP. Since our purpose is to illustrate how to replace the complicated bottom-up template matching strategy in B&L algorithm, we shall concentrate on testing the COP using PC-trees.

Our improvement over the B&L algorithm is based on a simple observation that there exist at most two special partial nodes and the unique path connecting them gives a streamline view of the update operation. In particular, no template is necessary in our PC-tree algorithm and the node-by-node examination is replaced by one swift batch operation.

2. A Forbidden Structure of Consistent Permutations

A key observation that simplifies our PC-tree algorithm is based on the following forbidden structure of consistent permutations.

**Theorem 1.** The following structure \( \{S_1, S_2, S_3\} \) is forbidden in any consistent permutation of the columns of \( M \) that preserves the COP.

Let \( S_1 = \{a_1, b_1\} \), \( S_2 = \{a_2, b_2\} \) and \( S_3 = \{a_3, b_3\} \) be three subsets of columns such that

1. Columns \( a_1, a_2 \) and \( a_3 \) are distinct from each other.
2. None of the \( a_i \)'s is the same as any of the \( b_i \)'s (but \( b_1, b_2 \) and \( b_3 \) are not necessarily distinct).
3. No two columns in any \( S_i \) are separated by any column in the other two subsets.

**Proof.** If \( b_1 = b_2 \), then from (3), we must have the arrangement \( a_1…b_1(=b_2)…a_2… \)

But then, it would be impossible to place \( b_3 \) anywhere in the arrangement unless \( b_1 = b_2 = b_3 \). In the latter case it would be impossible to place \( a_3 \) anywhere in the arrangement.

Hence, assume \( b_1 \neq b_2 \). From (3), we must have the following arrangement \( a_1…b_1…b_2…a_2… \) (or the reverse) in any consistent permutation of the columns. Since \( b_1, b_2 \) and \( b_3 \) must not be separated by the \( a_i \)'s, we shall have \( a_1…b_1…b_2…b_2…a_2… \). Note that in the above arrangement, \( b_3 \) is allowed to be the same as either \( b_1 \) or \( b_2 \). But then, it would be impossible to place \( a_3 \) anywhere in the arrangement.

3. Terminal Nodes and Paths

Define a partial node to be a terminal node if none of its children is partial. In other words, each child of a terminal node is either empty or full. Terminal nodes play a major role in simplifying our tree-update operation.

**Theorem 2.** If the given matrix \( M \) satisfies the COP, then there can be at most two terminal nodes in \( T \) at every iteration.
Proof. Suppose, on the contrary, there are three terminal nodes $t_1$, $t_2$, and $t_3$. Choose an empty child $a_i$ and a full child $b_i$ from each of the $t_i$, $i = 1, 2, 3$. Then, $S_1 = \{a_1,b_1\}$, $S_2 = \{a_2,b_2\}$ and $S_3 = \{a_3,b_3\}$ constitute a forbidden structure as described in Theorem 1.

If there are two terminal nodes $t_1$ and $t_2$, define the unique path connecting $t_1$ and $t_2$ to be the terminal path. In case there is only one terminal node $t$, the terminal path is defined to be the unique path from $t$ to $\text{ROOT}(T,u)$. In the latter case, $t$ is called a special terminal node if it is a C-node and its two endmost children are empty. That is, its full children are consecutively arranged in the middle.

Now, any C-node in the interior of the terminal path has its children (not on the path) divided into two sides. Since one can flip the children ordering along the terminal path to obtain an equivalent PC-tree, we shall show that there is a unique way to flip the children of these C-nodes “correctly”. Denote $\text{ROOT}(T,u)$ by $m$.

Theorem 3. If $M$ satisfies the COP and there are two terminal nodes $t_1$ and $t_2$, then
(1) Every C-node in the interior of the terminal path satisfies that its full children together with its two neighbors on the path are consecutive in its circular list. Therefore, they can be flipped correctly.
(2) Node $m$ must be on the terminal path.
(3) Let $w$ be a child of a node (other than $t$) on the unique path from $m$ to the root of $T$ such that $w$ itself is not on this path. Then $w$ must be empty.

Proof. Consider the following cases:
(1) If a C-node $w$ does not have any full child, we are done. Hence, assume $w$ has a full child $d$. Let the two neighbors of $w$ in the terminal path be $s_1$ and $s_2$. Suppose, in traversing the circular list of $w$ from $s_1$ to $s_2$ through the full child $d_1$, we encounter an empty child $d_2$, say in the order $s_1, d_1, d_2, s_2$. Let $d_1', d_2'$ be two leaves of $d_1, d_2$, respectively. Choose an empty child $a_i$ and a full child $b_i$ from each of the $t_i$, $i = 1, 2$.

Then $S_1 = \{a_1,b_1\}$, $S_2 = \{a_2,b_2\}$ and $S_3 = \{d_1',d_2'\}$ constitute a forbidden structure as described in Theorem 1.
(2) This is obvious.
(3) Suppose $w$ has a leaf $b$ in $u$. Since we have ...$a_1...b_1...b_2...a_2$...(or its reverse) in any consistent permutation of the columns, it would be impossible to place $b$ anywhere in the arrangement.

Theorem 4. Consider the case that $M$ satisfies the COP and there is only one special terminal node $t$. Let $w$ be a child of a node (other than $m$) on the path from $m$ to the root of $T$ such that $w$ itself is not on the path. Then $w$ must be empty.

Proof. Suppose $w$ has a leaf $b$ in $u$. Let $a_1, a_2$ be any leaf of the two endmost children of $t$, respectively and $b'$, a leaf in a full child of $t$. Then we shall have $a_1...b'...a_2$ (or its reverse) in any consistent permutation of the columns. But then, it would be impossible to place $b$ anywhere in the arrangement.

By considering the terminal nodes we have a global view of the distribution of all full nodes by Theorems 3 and 4. Figure 3 illustrates an example with two terminal
nodes. Note that, in this example, we have flipped all full children of nodes on the terminal path down.

![Diagram of a PC-tree](image)

**Fig. 3.** The unique separating path between two terminal nodes $u$ and $u'$

### 4. Constructing the New PC-Tree

Rather than using the node-by-node tree modification in the B&L algorithm, we shall update the current PC-tree to the new one in a batch fashion.

When a new row $u$ is added, perform the following labeling operation:
1. label all leaf columns in $u$ full.
2. the first time a node becomes partial or full, report this to its parent.
3. the first time a node $x$ gets a partial or full child label $x$ partial
4. the first time all children of a node $x$ become full label $x$ full

Our batch tree-update operation consists of the following steps:
1. delete all edges on the terminal path
2. duplicate the nodes on the terminal path (this is called the *splitting operation*)
3. create a new $C$-node $w$ and connect $w$ to all duplicated nodes according to their relative positions on the terminal path as follows (this is called the *modifying operation*): connect $w$ directly to all $P$-nodes; connect $w$ to all full children of $C$-nodes according to their original ordering; contract all degree two nodes

The following figures illustrate the splitting and the modifying operation on the graph of figure 3. Note that the root in a PC-tree does not play a major role in these operations.

![Diagram of the new PC-tree](image)

**Fig. 4.** The splitting operation
Theorem 5. The corresponding $PQ$-tree of the newly constructed $PC$-tree is the same as the one produced by the $B&L$ algorithm.

Proof. As noted before, the $B&L$ algorithm modifies the tree in a node-by-node bottom-up fashion based on 9 templates. Whenever a pattern is matched, it is substituted by the replacement.

We shall prove the theorem by induction on the depth of the pertinent subtree. For each template operation of the $PQ$-tree, we shall illustrate the corresponding splitting operation of the $PC$-tree and show the equivalence of the two operations in terms of the resulting $PQ$-tree. We shall skip the easy cases of $P_0$ and $P_1$, $Q_0$ and $Q_1$.

Fig. 5. Connecting to the new C-node

Fig. 6. The modifying operation

Fig. 7. The template operations of Booth and Lurker’s algorithm
First, consider templates at the root of $T$. In Figure 8 we consider template $P_2$. The top part of Figure shows the $PQ$-tree replacement. The bottom part gives the corresponding $PC$-tree splitting and modifying operation. The equivalence is shown by the two rightmost diagrams, in which a $PQ$-tree (on top) and its corresponding $PC$-tree (in the bottom) are obtained through $B&L$ template matching and the $PC$-tree operation, respectively.

The same goes for the templates $P_4$, $P_6$, $Q_2$ and $Q_3$ as shown in Figures 9, 10, 11, 12.

![Diagram](image-url)
Now, assume the theorem is true for $PQ$-trees whose pertinent subtree is of depth $k$ and consider a tree whose pertinent subtree is of depth $k+1$. In the $PQ$-tree operation, a template will be applied to a non-root terminal node $x$ of the current $PQ$-tree $T_1$, to obtain a new $PQ$-tree $T_2$. To simplify the argument we assume the pertinent subtree of $T_2$ has depth $k$. Since, by induction, further $PQ$-tree operations on $T_2$ can be done equivalently through the corresponding $PC$-tree operations, we only have to show that the template matching operation on node $x$ results in the same subtree as the one obtained by operation on the corresponding $PC$-tree of $T_2$. These are illustrated in Figures 13 and 14.

Fig. 10. Template P6 for $ROOT(T,S)$ when it is a doubly partial $P$-node

Fig. 11. Template Q2 for a singly partial $Q$-node

Fig. 12. Template Q3 for a double partial $Q$-node
Fig. 13. Template $P_3$ for a singly partial $P$-node which is not $ROOT(T,S)$

Fig. 14. Template $P_5$ for a singly partial $P$-node, other than $ROOT(T,S)$, with one partial child

End of Proof of Theorem 5

5. The Complexity Analysis

In the $B&L$ algorithm, the bottom-up propagation potentially could examine the same node many times. Hence, special care must be taken to ensure efficiency. In our splitting and modifying operation of $PC$-trees the update is done in one batch.

We shall apply the same pointer strategy as in the $B&L$ algorithm with regard to $P$-nodes and $C$-nodes. Our major saving is in skipping the pattern-matching phase at each node. The remaining replacement phase will be the same as in the $B&L$ algorithm except that we use $C$-nodes instead of $Q$-nodes.

The detailed operations needed for $PC$-trees are all included in those of the $B&L$ algorithm. Hence, the running time is linear in the size of the input.

6. Conclusion

We believe that a more efficient implementation of our $PC$-tree algorithm deviating completely from the pointer operation of $B&L$ algorithm is possible. However, the purpose of this paper is to demonstrate that, conceptually, there is a simpler way to view the $PQ$-trees and their updates. Implementation issues of $PC$-trees will be left for those who have the experience of implementing the $B&L$ algorithm.
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References

Abstract. We investigate the relative efficiency of a finite number of stacks in comparison to several variants of deques. In the nondeterministic setting, two stacks can simulate a general deque in linear time. This implies a negative answer to the question raised by Brandenburg whether a deque can simulate a finite number of tapes in linear time. We also show that in realtime an output-restricted deque cannot simulate two stacks for deterministic computations. It is known that a general deque can be simulated deterministically by three stacks in linear time. We describe an approach that is simpler to analyze and has a smaller constant factor (with respect to the required stack operations) than a previous solution.

1 Introduction

Functional programming with LISP-like data raises the question whether familiar data structures like queues and deques (double-ended queues) have efficient implementations in the available ‘one-ended’ lists or stacks. By efficient we mean that the simulator operates in (cumulative) linear time or realtime, i.e., the number of internal operations for each simulated step is bounded by a constant. Notice that the latter is stronger, since it rules out techniques like garbage-collection. On an abstract level, it has been known for a long time that sequential storage with multiple access-pointers—and hence queues and deques—can be simulated by sequential storages with single pointers in linear time [17] and even realtime [5,11]. Since each linear storage can be split into two stacks, this shows the somewhat surprising fact that (up to constant factors) no speed is lost if queues and deques are implemented in functional languages. Concrete implementations of queues have been given by Hood and Melville [8], while Chuang and Goldberg extended this to deques [4]. The known realtime simulations however require a large number of stacks per simulated queue or deque, while there is an obvious way to do a linear time simulation of a single queue by two stacks. It is also known that a deque can be simulated in linear time by three stacks [16]. Questions that arise are: What is the minimum number

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of stacks sufficient for linear or realtime simulation of queue and deque? Does the simulation of a deque become easier if the access to data is restricted?

In the area of formal languages, Ayers [1] investigated automata with a single deque that work in realtime (there called quasi-realtime). In his comments on [1], Brandenburg [3] stated several open problems. He asked whether the hierarchy of language-families accepted by nondeterministic machines working in realtime (or linear time) with a queue, an output-restricted deque, a deque, or a finite number of tapes is proper. These questions have partially been answered. Li e.a. [13] showed that in general even a deterministic machine with a single stack (which obviously can be simulated by an output-restricted deque) cannot be simulated in linear time by a nondeterministic machine with a single queue. An output-restricted deque can be simulated by two stacks in linear time, which by the nondeterministic realtime simulation of many tapes by three stacks [2] and the lower bound of [12, Corollary 3] shows that an output-restricted deque is weaker than a finite number of tapes.

The results presented here concentrate on a small number of stacks and the relation to variants of the deque.

2 Preliminaries

We adopt the terminology from [10]. In our investigations, stacks, queues, and deques are linear lists of atomic items or symbols. Information can be stored and retrieved (along with each deletion) at the ends of the list. In addition, emptiness of the storage can be sensed. More precisely we distinguish between:

Stack: All insertions (push operations) and deletions (pop operations) are made at one end.
Queue: All insertions are made at one end, all deletions at the other end.
Output-restricted deque: Insertions can be made at both ends of the list, while deletions are allowed at one end only.
Input-restricted deque: Deletions can be made at both ends of the list, while insertions are allowed at one end only.
Deque: Insertions and deletions can be made at both ends of the list.

Note that in comparison to [10] we exclude any access to the interior of the lists.\footnote{What is called stack here is also known as a pushdown-store. The name ‘stack’ is often used for a storage that admits read access to its interior, which we don’t allow.}

The usual framework in which these data storages are compared is by attaching them to a finite control that has a separate read-only one-way input-tape and either accepts or rejects an input, see e.g. [15,13].

3 Nondeterministic Simulation of a Deque by Two Stacks

We first consider nondeterministic machines and show that—up to a linear time overhead—machines with a single deque are equivalent in accepting power to those equipped with two stacks.
Theorem 1. Every \( t(n) \) time-bounded nondeterministic deque machine can be simulated by an \( O(t(n)) \) time-bounded nondeterministic two-stack machine.

Proof. We will divide the simulation into two stages. The first stage simulates in realtime a deque by a queue and two stacks \( S_L \) and \( S_R \) with some restrictions on the way the queue can be accessed. In the second stage we show how to simulate a queue and two stacks with two stacks within the time-bound claimed in the theorem.

Without loss of generality we assume that the deque is empty when the computation terminates. We claim that a single deque can be simulated by a queue and two stacks step by step as follows. The simulator starts with empty stacks and an empty queue. It reads input and executes operations in accordance to the finite control of the deque machine. In addition the simulator is in one of two modes, either in left-right or in right-left mode (these names will be explained later). The initial mode can be chosen arbitrarily and during the simulation the simulator may switch from one mode to the other whenever the queue is empty. The possible deque operations are now simulated in the following way:

**Insert at the left end:** Nondeterministically choose to push the new symbol onto \( S_L \) or, provided \( S_L \) is empty and the current mode is left-right, insert the new symbol into the queue.

**Insert at the right end:** Nondeterministically choose either to push the new symbol onto \( S_R \) or, provided \( S_R \) is empty and the current mode is right-left, insert the new symbol into the queue.

**Delete at the left end:** If \( S_L \) is not empty, pop one symbol. If \( S_L \) is empty, remove one symbol from the queue, provided it is not empty and the current mode is right-left. Otherwise abort the simulation.

**Delete at the right end:** If \( S_R \) is not empty, pop one symbol. If \( S_R \) is empty, remove one symbol from the queue, provided it is not empty and the current mode is left-right. Otherwise abort the simulation.

**Sense emptiness:** Check that \( S_L \), \( S_R \), and the queue are empty.

The main invariant of the simulation (which holds after the simulation of one deque access) is, that the current deque contents of the simulated machine are always encoded by the concatenation of the contents of \( S_L \) with its top to the left, the queue either with its rear to the left (in left-right mode) or to the right (in right-left mode), and \( S_R \) with its top to the right. By checking the simulations of the four operations on the deque and their subcases it can be verified that this invariant is maintained and that the access is always to the symbols at the encoding of the ends of the deque. We thus see that each successful run of the simulator corresponds to an accepting computation of the deque machine.

Conversely, we have to argue that every accepting computation of the deque machine can be simulated successfully. Fix an accepting computation of the deque machine. We can associate with every symbol stored on the deque (which we assume to be distinguishable for our considerations) two bits of information: whether it entered the deque from the left or the right end and whether it will leave the deque at the left or the right end (we assumed that the deque
will eventually be empty). The symbols will be sorted according to each of the components, which leads to the important observation that in each configuration of the deque there is at most one type of symbols with a ‘mixed mode’ (left-right or right-left). Let us assign to each configuration of the deque the ‘mixed mode’ of its contents (or an arbitrary mode if there is no symbol with mixed mode). The simulator now follows the information associated with the symbols and always works in the mode prescribed by the mode of the configuration of the deque. It does so by storing the left-left symbols on $S_L$, the right-right symbols on $S_R$, and all mixed-mode symbols (left-right or right-left) on the queue. Clearly the mode of the deque can only change when there is no mixed-mode symbol, and then the queue of the simulator is empty allowing it to switch between the modes as well. If the deque contains symbols that should be deleted at the left end, the simulator will either store them on $S_L$, where they are readily available, or they are mixed-mode symbols, which forces them to be at the front of the queue where they can be accessed by the simulator. An analogous argument shows that deletions at the right end can be carried out faithfully.

By the above discussion each computation of the deque machine can be simulated and each run of the simulator corresponds to a computation of the deque machine.

Now we show how to simulate the two stacks and one queue on two stacks. The simulated stacks will be mapped onto the corresponding stacks of the simulator, with the additional property that there is a bottom marker below each portion encoding a simulated stack. Below the bottom markers two strings are stored, where the shorter string is a suffix of the longer if the simulation is to be completed successfully (if both strings have the same length, they should be identical). If the simulated machine is in left-right mode, then the string below the marker stored on $S_L$ is at least as long as the one on $S_R$ and vice versa for right-left mode. If a symbol is to be entered into the queue, then this symbol is added to the ‘longer’ string, if a symbol is deleted it is guessed and added to the shorter string. Note that accessing the strings below the marker is always possible since the corresponding stack will be empty at that point of the simulation.

If eventually an accepting state has been reached, the contents of the two stacks are compared symbol by symbol in $O(t(n))$ time and the input is accepted if they are identical.

The simulation strategy described so far admits a simulation of each valid computation. It may however include invalid computations as well for two reasons. First, switching from one mode to the other is allowed only if the queue is empty, and the simulator has to verify this condition. Second, even if switching from one mode to the other occurs only at valid points during the simulation, symbols might be deleted from the simulated queue before they have been inserted. The first problem is overcome by inserting a special zero marker into the guessed strings after the last symbol has been entered resp. deleted just before switching the mode. Consider a situation when the simulator is in left-right mode. Then after the last symbol during this phase has been stored on $S_R$ the
simulator adds the zero marker to the string on $S_R$ and records that the marker has been set. Note that above this zero marker the stack can still be simulated, but further insertions are not allowed until another zero marker is stored on $S_L$ and the mode changes to right-left. The simulator works analogously in right-left mode. Zero markers necessarily occur in pairs, one on each stack of the simulator. Therefore the length of the strings at steps when the mode changes are equal (assuming that the final comparison is successful) and the simulated queue is empty at these points.

For the second problem we associate with each symbol a flag from the set \{0, 1\} and we let the simulator remember the current flag $f$. With each symbol inserted into the queue, the simulator associates $f + 1 \pmod{2}$, with every deleted symbol the flag $f$. If the current value of the flag has been used for at least one insertion it may nondeterministically be incremented by 1 modulo 2. Assume for a contradiction that the resulting strings could be equal although at least one symbol has been deleted from the queue being simulated before it was inserted. Consider the situation just before this happens for the first time. The lengths of the strings will be identical and the flags of the topmost pair of symbols match, both having the value $f$ (since we assumed that the strings would eventually be equal). The flag has been incremented since the last insertion without having been used for another one. No further modification of the flag is possible before the next symbol is inserted into the simulated queue. Therefore the (erroneously) deleted symbol gets flag $f$ as well. The next symbol to be inserted will have flag $f + 1 \pmod{2}$, which will cause a mismatch, contradicting our assumption.

Since each step in the two stages of the simulation can be carried out in constant time, we see that the entire simulation can be completed in $O(t(n))$ steps.

Together with the obvious simulation of two stacks by a deque (the stacks are stored with bottom markers next to each others at both ends of the deque) we obtain:

**Corollary 1.** For nondeterministic machines a single deque and two stacks are equivalent up to a linear time overhead.

We can now answer the question from [3] whether nondeterministic machines with several tapes working in linear time are stronger than linear-time machines with one deque. Informally we obtain that two tapes (which are equivalent to many tapes [2]) are stronger than a single deque.

**Corollary 2.** There is no general linear time simulation of multiple tapes by one deque for nondeterministic machines.

Proof. Suppose that a linear time simulation of two tapes by a deque exists. For the special case of realtime acceptance by two tapes we obtain an upper bound of $O(n)$ on two stacks. Simulating these two stacks by the method of [12] on a single tape we obtain an $O(n^{1.5} \sqrt{\log n})$ upper bound on the time. But this contradicts the known almost quadratic lower bounds for the nondeterministic simulation of two tapes by one tape, e.g. $\Omega(n^2/\log^k n)$ for any positive $k$ from [6] ($\log^k$ denotes the $k$-fold iteration of log).
By calculating the simulation overhead we can even derive the lower bound $\Omega(n^{4/3}/\sqrt{\log n})$ for the nondeterministic simulation of two tapes by one deque.

For nondeterministic computations three stacks are as powerful as any finite number of tapes \cite{1}. Simulations by three or more stacks are therefore interesting only in the deterministic setting. In the next section we will consider some simulations of this kind.

4 Deterministic Simulations

For deterministic simulations of a deque by two stacks we have a less complete picture than in the nondeterministic setting. The well-known linear time simulation of a queue by two stacks (see, e.g., \cite{3,13} or \cite{10, Ex. 2.2.1.14}) generalizes to output-restricted deques, while we don’t have an analogous result for input-restricted or general deques.

The question whether there is a real-time equivalence between two stacks and an output-restricted deque is answered negatively by the following results.

Let $u^R$ denote the reversal of string $u$. Define $L \subseteq \{0, 1, 2, 3, \$\}^*$ as

$$L = \{w_1\$w_2\$w_3xy\$z \mid w_1, w_2, w_3 \in \{0, 1\}^*, x \in \{2, 3\}, y = w_2^R \text{ if } x = 2, \ y = w_3^R \text{ if } x = 3, \ \exists t \in \{0, 1\}^*: w_1 = tz^R\}.$$

Lemma 1. There is a deterministic two-stack machine that accepts $L$ in real-time.

Proof. Clearly, the machine can decide membership in the regular language \{0, 1\}*$\{0, 1\}*$\{0, 1\}*$\{2, 3\}*$\{0, 1\}*$ on the fly. We may therefore assume that the input can be decomposed according to the definition of $L$. The machine accepting $L$ pushes copies of $w_1$ on both of its stacks and marks these segments of its stacks. Then it copies $w_2$ onto stack 1 and $w_3$ onto stack 2. When $x$ is encountered it accordingly compares the string stored on stack 1 or stack 2 with the input segment $y$ following $x$. If $y$ matches, the remaining symbols of the input after the last $\$ are compared with the topmost section of the stack which has just been emptied (discarding the marker first).

Lemma 2. Deterministic machines with one output-restricted deque cannot accept $L$ in real-time.

Proof. We assume that $L$ can be accepted by a machine $M$ of the kind described in the lemma (with a deque that allows deletions only at its rear) and derive a contradiction. We simplify the proof by assuming that $M$ executes one step per input symbol, which can be achieved by possibly increasing the number of internal states and of storage symbols.

The proof uses the notion of incompressibility, in particular the non-constructive existence of incompressible binary strings of each length, see \cite{14}. We will also require some strings of symbols to be self delimiting, which means that
their last symbols can be detected even if they are concatenated with other strings. A simple (though not very efficient) way to achieve this is to repeat each bit except for the last one, which is encoded by one of 01 or 10.

Let \( m \) be the number of different symbols available to \( M \) for storing information on its deque. Fix the constant \( c = \max(2, 3\lceil \log_2 m \rceil) \). Choose an incompressible string \( w \) of length \((c^2 + 2c + 1)n\) for \( n \) sufficiently large to make the argument work. Divide \( w \) into three portions \( w = w_1w_2w_3 \) with \(|w_1| = (c^2 + c)n\), \(|w_2| = n\), and \(|w_3| = cn\). We now consider accepting computations of \( M \) on input strings with a prefix \( w_1\$ \). Our first claim is that after having read this prefix the length \( \ell \) of \( M \)'s deque exceeds \( 2(c + 1)n \). We can recover \( w_1 \) from the contents of the deque by feeding the string \( \$2\$ \) into \( M \) (this will encode three empty strings \( w_2, w_3 \), and \( y \)) and then looking symbol by symbol for the longest trailing string that causes \( M \) to accept. Since any string that leads to acceptance is a prefix of \( w_1^R \), this will recover \( w_1^R \). Now suppose \( \ell \leq 2(c + 1)n \). The entire string \( w \) can be described by:

- A self delimiting description of \( M \) of size \( O(1) \).
- The state of \( M \) after reading \( w_1\$ \) of size \( O(1) \).
- A self delimiting description of the above algorithm for recovering \( w_1 \) of size \( O(1) \).
- A self delimiting encoding of \( \ell \) in \( O(\log n) \) bits.
- A self delimiting encoding of \( n \) in \( O(\log n) \) bits.
- The contents of \( M \)'s deque encoded in binary in at most \( \ell\lceil \log_2 m \rceil \leq \frac{2}{3}(c^2 + c)n \) bits.
- The strings \( w_2 \) and \( w_3 \) (no encoding is necessary, since their lengths are known).

For sufficiently large \( n \) this will result in a compression of \( w \) contradicting its choice.

Notice that after reading \( w_1\$ \) the deque is now longer than the remaining suffix of any accepted input of the form we consider with an empty trailing portion \( z \). This implies that the first symbol on \( M \)'s deque at this point will never be removed and nothing added at the head can later influence the computation.

Consider next the state of \( M \)'s deque when \( w_1\$w_2\$w_3 \) has been read. We count the number of trailing positions of these deque contents that have been modified while \( M \) was reading \( w_3 \). Assume first that this number exceeds \( n + 1 \). Intuitively, \( M \) has—up to a finite number of bits—hidden (or even destroyed) the encoding of \( w_2 \). If the trailing \( n + 2 \) symbols on the deque are known, then \( w_2 \) can be recovered by starting \( M \) in its state \( q_1 \) after reading \( w_1\$w_2\$w_3 \) and feeding in strings of the form \( 2u\$ \) with \( u \in \{0, 1\}^n \) until the unique string \( 2w_2^R\$ \) leading to acceptance has been found. The trailing segment of length \( n + 2 \) of the deque can be generated by starting \( M \) from the state \( q_2 \) it was in before it accessed the first symbol in this segment for the last time and feeding in the portion of \( w_3 \) that had not been processed at this point.

We can now describe \( w \) by:

- A self delimiting description of \( M \) of size \( O(1) \).
- The states $q_1$ and $q_2$ in $O(1)$ bits.
- A self delimiting description of the above algorithm for recovering $w_2$ of size $O(1)$.
- A self delimiting encoding of $n$ in $O(\log n)$ bits.
- A self delimiting encoding of the position within $w_3$ from which $M$ has to be started in order to generate the trailing portion of the deque in $O(\log n)$ bits.
- The strings $w_1$ and $w_3$ (again, their length is available).

For $n$ sufficiently large we obtain a compression contradicting the choice of $w$.

The other case is where a trailing string $u$ of at most $n + 1$ symbols of the deque includes all positions modified while reading $w_3$. It is dealt with in a similar way, now compressing $w_3$. \hfill \square

As a consequence of the two preceding lemmas we obtain:

**Theorem 2.** Deterministic machines with one output-restricted deque cannot simulate two-stack machines (and hence deque machines) in realtime.

Two stacks cannot recognize the language $\{x2x' \mid x \in \{0, 1\}^* \text{ and } x' \text{ is a prefix of } x\}$ by the (stronger) result of [9] in realtime, while even with a queue this language can be accepted. Therefore two stacks and an output-restricted queue are of incomparable power for deterministic realtime computations.

Next we consider simulations by more than two stacks. Rosenberg [16] describes a linear time simulation of a general deque by three stacks. The idea behind the simulation is again to maintain a left and a right section of the deque on two stacks. Whenever exactly one of these sections is empty and a delete operation on the respective portion is requested, the third stack is used in order to partition the non-empty section into segments of approximately half of its original size. The analysis of this simulation is somewhat difficult, since items may be copied from one section to the other and back an unbounded number of times.

We present here a different linear time simulation of a deque by three stacks, which is easier to analyze and has a smaller linear factor in terms of the number of push and pop operations required. While Rosenberg’s solution uses at most $18n$ stack-operations for simulating $n$ deque-operations, our bound will be $9n$.

We remark that Rosenberg’s simulation is closely related to the construction used by Stoss [17] in the simulation of multi-head tape-units by single-head ones, while ours is inspired by the very first proof of such a simulation given by Hartmanis and Stearns [7].

**Theorem 3.** A sequence of $n$ deque operations can be simulated deterministically by three stacks with $9n$ stack-operations.

Proof. Like in Rosenberg’s simulation, we represent a deque as two stacks $S_L$ and $S_R$ that encode the current contents of the deque. In contrast to the simulation outlined above, where (except for intermediate states) the concatenation of the two stacks reflects the deque contents verbatim, in our simulation some information may be duplicated or outdated. Therefore some deletions at the bottom
of the stacks might be necessary in order to obtain the current contents of the simulated deque. The membership of an item in the section possibly outdated or duplicated will be expressed by coloring it red, while all other items will be green. These colors can either be represented by a single bit attached to each item, or by markers on the stacks, since items of the same color form blocks. Initially $S_L$ and $S_R$ are empty, as is the third stack $S_T$ that will hold a counter which has the value 0 at the start of the simulation. Insert operations are simply mapped to push operations of green items onto $S_L$ or $S_R$. We will now describe a delete operation on the left in detail. The corresponding operation on the right works symmetrically.

If $S_L$ contains an item that is colored green as its topmost item, then it is returned. If $S_L$ has a red item as its topmost and the count maintained on the third stack is not 0, then the red item is returned and the counter is decremented. Otherwise a clean-up phase is initiated. All remaining items on $S_L$ are deleted. The green items on top of $S_R$ are popped one after the other and pushed onto $S_L$ as well as onto $S_T$, changing their color to red. The remaining red items on $S_R$ are discarded. The items now on $S_T$ are moved back onto $S_R$ and their number is counted on $S_L$ above its red items. Finally the count is transferred to $S_T$. Then the simulation is resumed and the deque operation is carried out, if possible.

In order to see that the simulation is correct, note that the current contents of the deque being simulated is encoded by the green items on $S_L$ (top to bottom), a top segment of $c$ red items on $S_L$ (where $c$ is the value currently stored on $S_T$), or equivalently the mirror image of a top segment of red items on $S_R$, and finally the green items on $S_R$ from bottom to top. This invariant is clearly maintained by push operations and pop operations of green items. A pop operation of a red item reduces the count and thus adapts the length of the red segments considered on both stacks. The clean-up moves a sequence of green items into fully overlapping red segments.

Each item enters in green color, causing one push (1 operation). If it leaves the deque while still having green color, the total cost is 2 operations. Otherwise it changes its color to red. While doing so, it is moved from $S_L$ or $S_R$ to the other stacks (3 operations), one of the copies is moved back from $S_T$ and counted (3 operations), and the count is transferred to $S_T$ (2 operations). If it is never removed, then both copies stay in the stacks and the total cost is 9 operations. Otherwise the item is removed and possibly also its other copy (2 operations). In addition the count is updated (1 operation). In the latter case the total is at most 12 push or pop operations for simulating two operations, which yields 6 operations per simulated deque operation.

In the simulation given in the previous proof red items can be left on the stacks in the clean-up phase (which then doesn’t really clean-up), saving $n$ operations at the cost of wasting stack space. Also, the counter values could be reduced by a linear factor, which would save $(3 - \epsilon)n$ operations for any $\epsilon > 0$, an improvement which can be applied to Rosenberg’s simulation as well.
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References

Optimizing a Computational Method for Length
Lower Bounds for Reflecting Sequences

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Abstract. We refine and optimize a computationally intensive enumeration method, based on the traversal of a quadtree, for finding lower bounds on the lengths of reflecting sequences for labeled chains. The improvement results from the introduction of a redundancy relation defined on vertex-pairs of the underlying quadtree, which enables the pruning of redundant branches near the root of the quadtree, as well as locally at deeper depths. The test run of the implementation showed a length lower bound of $19t - 214$ for $t$-reflecting sequences for labeled 7-chains with significant speedup, which yields the current length lower bound $\Omega(\frac{n^{1.51}}{n^{17} + 1})$ for universal traversal sequences for 2-regular graphs of $n$ vertices, and $\Omega(d^{2 - 1.51} n^{2.51})$ for universal traversal sequences for $d$-regular graphs of $n$ vertices, where $3 \leq d \leq \frac{n}{17} + 1$.

1 Preliminaries

Reflecting sequences are variants of universal traversal sequences, and were introduced by Tompa in proving lower bounds on the lengths of universal traversal sequences. The study of universal traversal sequences is motivated by the complexity of graph traversal. Good bounds on the lengths of universal traversal sequences translate into good bounds on the time complexity of certain undirected graph traversal algorithms running in very limited space.

Aleliunas et al. proved the existence of polynomial-length universal traversal sequences, yielding polynomial-time and logarithmic-space (non-uniform) deterministic algorithms for undirected graph traversal. Beame et al. used variants of the jumping automaton for graphs of Cook and Rackoff to study time-space tradeoffs for undirected graph traversal. Good lower bounds on the length of universal traversal sequences provide a prerequisite to proving time-space tradeoffs for traversing undirected graphs for these computational models.

1.1 Universal Traversal Sequences for Undirected Graphs

For a graph $G$, denote by $V(G)$ the vertex set of $G$, and by $E(G)$ the edge set of $G$. For positive integers $d$ and $n$ such that $d < n$, let $\mathcal{G}(d, n)$ be the set of all connected $d$-regular graphs of order $n$ with an edge-labeling. The edge-labeling on $G \in \mathcal{G}(d, n)$ is defined as follows: For every edge $\{u, v\} \in E(G)$, there are two
labels \( l_{u,v}, l_{v,u} \in \{0,1,\ldots,d-1\} \) with the property that for every \( u \in V(G) \), \( \{l_{u,v} \mid \{u,v\} \in E(G)\} = \{0,1,\ldots,d-1\} \). Note that \( G(d,n) \) is not empty if and only if \( dn \) is even. A sequence \( U \in \{0,1,\ldots,d-1\}^* \) can be considered as a sequence of edge-traversal commands and induces a walk starting at any vertex in every labeled graph in \( G(d,n) \). We say that such a sequence \( U \) traverses \( G \in G(d,n) \) starting at \( v_0 \in V(G) \) if, by starting at \( v_0 \) and following the sequence of edge labels in \( U \), all the vertices in \( G \) are eventually visited. The sequence \( U \) is called a universal traversal sequence (UTS) for \( G(d,n) \) if \( U \) traverses every \( G \in G(d,n) \) starting at every vertex in \( G \).

Let \( U(d,n) \) denote the length of a shortest UTS for non-empty \( G(d,n) \), and define \( U(d,n) = U(d,n+1) \) in case \( G(d,n) \) is empty. The lower and upper bounds on \( U(d,n) \) for various ranges of \( d \) were studied in [1], [3], [4], [7], [9], [5], and [8]. Currently, the best lower bounds on \( U(d,n) \) are in [8],

\[
U(d,n) = \begin{cases} \Omega(n^{\log_2 19}) & \text{if } d = 2, \\ \Omega(d^{2 - \log_2 19} n^{1 + \log_2 19}) & \text{if } 3 \leq d \leq \frac{n}{17} + 1. \end{cases}
\]

which were obtained using a computationally intensive approach.

### 1.2 Reflecting Sequences for Labeled Chains

For positive integer \( n \), a labeled chain of length \( n \) is a graph \( G \) with vertex set \( V(G) = \{0,1,\ldots,n\} \) and edge set \( E(G) = \{(i,i+1) \mid 0 \leq i < n\} \) with an edge-labeling defined as follows: Every edge \( \{(i,i+1) \in E(G) \) is associated with two labels \( l_{i,i+1} \) and \( l_{i+1,i} \), each a non-empty subset of \( \{0,1\} \), such that

1. \( l_{0,1} = l_{n,n-1} = \{0,1\} \), and
2. \( l_{i,i-1} \) and \( l_{i+1,i} \) form a partition of \( \{0,1\} \) for all \( 0 < i < n \).

Denote by \( L(n) \) the set of all labeled chains of length \( n \). A labeled chain \( G \in L(n) \) can be identified with the sequence \( \alpha = \alpha_1 \alpha_2 \cdots \alpha_{n-1} \in \{0,1\}^{n-1} \), where \( l_{i,i+1} = \{\alpha_i\} \) for \( 0 < i < n \).

Given a labeled chain \( G \in L(n) \), every sequence \( U = U_1 U_2 \cdots U_k \in \{0,1\}^k \) where \( k \geq 0 \), when considered as an edge-traversal sequence starting at vertex \( 0 \) in \( G \), determines a unique sequence \( 0 = v_0, v_1, v_2, \ldots, v_k \in \{0,1,\ldots,n\}^{k+1} \) such that \( U_i \in l_{v_{i-1},v_i} \) for all \( 1 \leq i \leq k \). For positive integer \( t \), a sequence \( U \in \{0,1\}^* \) is said to reflect \( t \) times in \( G \in L(n) \) if the walk in \( G \) induced by \( U \) alternately visits the end-vertices \( n \) and 0 at least \( t \) times; that is, there exist \( 0 < i_1 < i_2 < \cdots < i_t \leq |U| \) such that \( v_{i_{2j-1}} = n \) for all \( 1 \leq j \leq \lceil \frac{t}{2} \rceil \) and \( v_{i_{2j}} = 0 \) for all \( 1 \leq j \leq \lfloor \frac{t}{2} \rfloor \). The sequence \( U \) is called a t-reflecting sequence for \( L(n) \) if \( U \) reflects \( t \) times in every \( G \in L(n) \). Denote by \( R(t,n) \) the length of a shortest \( t \)-reflecting sequence for \( L(n) \).

The following theorem reduces the length lower bounds for universal traversal sequences to those for \( t \)-reflecting sequences.

**Theorem 1.** [3] Suppose that there exist positive integer \( c \geq 2 \) and positive reals \( r \geq 2 \) and \( k \) such that, for every positive integer \( t \), \( R(t,c) \geq rt - k \). Then

\[
U(d,n) = \begin{cases} \Omega(n^{\log_c r}) & \text{if } d = 2, \\ \Omega(d^{2-\log_c r} n^{1+\log_c r}) & \text{if } 3 \leq d \leq \frac{n}{17} + 1. \end{cases}
\]
In the light of Theorem 1, we can obtain length lower bounds for universal traversal sequences by seeking constants $c$, $r$, and $k$ such that every $t$-reflecting sequence for $L(c)$ has length at least $rt - k$, in which a large $\log_c r$ translates into a good length lower bound. The lemma below shows that it suffices to consider only the case when $t$ is odd.

**Lemma 1.** For all positive integers $t$ and $c$ and positive reals $r$ and $k$, if $R(2t - 1, c) \geq 2rt - k$ for every positive integer $t$, then $R(t', c) \geq rt' - k$ for every positive integer $t'$.

### 2 The Quadtree Enumeration Method

To show that $R(2t - 1, c) \geq 2rt - k$ in Lemma 1, we assign a minimum of $2^{c-1}t$ “marks” to various positions in a hypothetical $(2t - 1)$-reflecting sequence $S$ ($t$ marks for each labeled chain of length $c$). We argue that the “marks-to-bits” density for all but a short suffix of $S$ must be at most $2^{c-1}/2r$, implying a minimum length of $2^{c-1}t/(2^{c-1}/2r) = 2rt$ for all but a short suffix of $S$ having a bounded number of marks. We then determine $k$, depending on the maximum number of marks in the suffix, which gives that $R(2t - 1, c) \geq 2rt - k$.

Following the framework developed and analyzed in [8], it suffices to construct a tree $T$ that can be used to divide every marked $S$ into even-length segments, such that all but possibly the last segment has a marks-to-pairs density at most $2^{c-1}/r$. Specifically, $T$ is a quadtree whose vertices represent potential even bit-positions in a hypothetical $(2t - 1)$-reflecting sequence for $L(c)$, and whose branches from a position are labeled by the four possible pair-continuations at that position, that is, 00, 01, 10 and 11.

We summarize the Quadtree Enumeration method [8] in the following algorithm.

**algorithm** Build_Quadtree $(c, \rho)$

1. Initialize $S$ to the pair 00;
2. (* Only examine 00-branch by symmetry. *)
   While $S \neq 01$
   2.1. Try to find a prefix (pair-sequence) of $S$ whose marks-to-pairs density is at most $\rho$;
   2.2. If no such prefix exists, then
      2.2.1. Extend the quadtree by appending the pair 00 to $S$;
   else
      2.2.2. (* The sequence $S$ represents a leaf. *)
      Replace $S$ with the pair-sequence no longer than $S$ and following $S$ in the lexicographic ordering of pair-sequences;

**end** Build_Quadtree

To determine in step 2.1 an upper bound on the marks-to-pairs density of a sequence $S$, we assume the maximum number of marks possible for each chain in $L(c)$, which can be obtained by assuming the traversal of the chain starting at vertex 1 with the traversal sequence $S$. Hence, for each $C_i \in L(c)$, where $i \in \{0^{c-1}, 0^{c-2}1, \ldots, 1^{c-1}\}$, we traverse $C_i$ starting at vertex 1 with $S$, and place two types of marks $M_i$ on the pairs of $S$ as follows.

1. **Closed marks** of $M_i$ on $S$ that correspond to the last exit from vertex 1 during every traversal from vertex 0 to vertex $c$ with $S$, and
2. An open mark of $M_i$ on $S$ that corresponds to the last exit from vertex 1 during a traversal from vertex $c$ with $S$. Equivalently, the mark $M_i$ on $S$ is not closed on $S$ but it becomes closed on $SS'$ for some sequence $S' \in \{0,1\}^*$. We compute the desired upper bound in a graph-theoretic setting. Informally, for a marked traversal sequence $S$ above, we construct an inconsistency graph $G_c(S)$ of $S$, whose vertex set represents all possible open marks on $S$ and edge set represents all pairwise inconsistencies between open marks on $S$.

The algorithm for deciding if a particular $S$ constitutes a leaf in the quadtree is summarized below.

\begin{algorithm}
\textbf{Is\_Leaf} ($S$, $c$, $\rho$) return boolean
\begin{enumerate}
\item For each $i \in \{0^{c-1},0^{c-2},\ldots,1^{c-1}\}$, place closed and open marks of $M_i$ on $S$;
\item For each non-empty prefix (pair-sequence) $S'$ of $S$,
\begin{enumerate}
\item Form the inconsistency graph $G_c(S')$ of open marks on $S$, and solve the Maximum Independent Set problem for $G_c(S')$;
\item Compute the sum $m$ of the number of closed marks on $S'$ (in step 1) and the independence number of $G_c(S')$ (in step 2.1);
\end{enumerate}
\item If $m/\sqrt{|S'|} \leq \rho$, then return true;
end Is\_Leaf
\end{enumerate}
\end{algorithm}

A successful termination of Build\_Quadtree $\left(c, \frac{2^{c-1}}{r}\right)$ gives the supposition in Lemma 1.

**Theorem 2.** If Build\_Quadtree $\left(c, \frac{2^{c-1}}{r}\right)$ successfully terminates for some positive integer $c \geq 2$ and positive real $r \geq 2$, then the supposition (hence the conclusion) is achieved.

### 3 Refinements and Optimizations

In step 2 of the algorithm Is\_Leaf, each upper-bound computation for the marks-to-pairs density of every non-empty prefix $S'$ of segments $S$ of a hypothetical reflecting sequence induces an instance of the Maximum Independent Set problem for $G_c(S')$. We can encompass all these instances with a unified approach: each $G_c(S')$ is an induced subgraph of the same inconsistency graph $G_c$, which captures all (pairwise) inconsistencies between all possible pairs of open marks on a traversal sequence.

For a positive integer $c$, the inconsistency graph $G_c$ is defined as follows. The vertex set of $G_c$ is $\{(C_\alpha, u) \mid \alpha \in \{0^{c-1},0^{c-2},\ldots,1^{c-1}\} \text{ and } u \in \{2,3,\ldots,c-1\} \text{ is odd}\}$, in which a vertex $(C_\alpha, u)$ represents an open mark $M_\alpha$ on a non-empty traversal pair-sequence with the terminating vertex $u$ (hence $1 < u < c$ and $u$ is odd). The edge set of $G_c$ represents pairwise inconsistencies between all possible pairs of $((C_\alpha, u), (C_\beta, v)) \in V(G_c) \times V(G_c)$ from one trivial and two non-trivial sources.

The trivial source of inconsistency results from that each traversal pair-sequence yields at most one open mark $M_\alpha$ for each $\alpha \in \{0^{c-1},0^{c-2},\ldots,1^{c-1}\}$, so we have $((C_\alpha, u),(C_\alpha, v)) \in E(G_c)$ for all $(C_\alpha, u),(C_\alpha, v) \in V(G_c)$ with $u \neq v$. The two non-trivial sources are suffix-inconsistency and prefix-inconsistency. Informally, the suffix-inconsistency reflects the co-existence of two open marks,
while the prefix-inconsistency represents the co-reachability of the open marks from their initial configuration.

For $\alpha \in \{0,1\}^*$, i.e., $\alpha = \alpha_1 \alpha_2 \cdots \alpha_n$ for some $n \geq 0$, where $\alpha_i \in \{0,1\}$ for $i = 0,1,\ldots,n$, let $\alpha(i,j)$ denote $\alpha_i \alpha_{i+1} \cdots \alpha_j$ if $i \leq j$, and the empty sequence otherwise. We denote by $\overline{\alpha}$ the reversal of the component-wise complement of $\alpha$, that is, $\overline{\alpha} = \overline{\alpha_n} \cdots \overline{\alpha_1}$. To study the suffix-consistency of $(C_\alpha, u)$ and $(C_\beta, v)$ for two labeled chains $C_\alpha, C_\beta \in \mathcal{L}(c)$ and two vertices $u, v \in \{1,2,\ldots,c\}$ in $C_\alpha$ and $C_\beta$, respectively, we consider $\delta(u,v) = (c-u)-(v-1)$, which measures the difference of distances of $u$ and $v$ from vertices $c$ and 1 in $C_\alpha$ and $C_\beta$, respectively.

The theorem below characterizes the suffix-inconsistency between $(C_\alpha, u)$, $(C_\beta, v) \in \mathcal{L}(c) \times \{2,3,\ldots,c-1\}$ via $\delta(u,v)$ in the number of non-reflecting bits in $\alpha(2, c-1 - \delta(u,v))$ versus $\beta(2, c-1 - \delta(u,v))$. We define a first-order predicate $NR_{c,\alpha,\beta}$ for the maximum number of non-reflecting bits in such $\alpha, \beta \in \{0,1\}^{c-1}$ inductively as follows:

$$NR_{c,\alpha,\beta}(i) = \begin{cases} \beta(2, c-1 - i) = \alpha(2, c-1 - i) & \text{if } i \text{ is either 0 or 1 ,} \\
(\beta(2, c-1 - i) = \alpha(2, c-1 - i)) \lor \\
\exists j_i \in \{2,3,\ldots,c-1-i\} (\beta_{j_i} = \alpha_{2+c-i-j_i}) & \text{if } i \geq 2 . \end{cases}$$

**Theorem 3.** Let $(C_\alpha, u), (C_\beta, v) \in \mathcal{L}(c) \times \{2,3,\ldots,c-1\}$, we have:

1. If $\delta(u,v) < 0$, then $(C_\alpha, u)$ and $(C_\beta, v)$ are suffix-consistent.
2. If $\delta(u,v) \geq 0$, then $(C_\alpha, u)$ and $(C_\beta, v)$ are suffix-inconsistent if and only if $NR_{c,\alpha,\beta}(\delta(u,v))$.

### 3.1 Pruning Redundant Branches in the Quadtree

Every labeled chain $C \in \mathcal{L}(c)$ has an “ideal walk” starting at vertex 1 that represents the densest packing of reflections (marks) that $C$ can exhibit; and that sequence of vertices visited depends only on $c,1,2,\ldots,c-1,c-2,\ldots,1,0,1,\ldots$. We can measure the goodness of every walk in a labeled chain $C$ starting at vertex 1 by how much of the ideal walk in $C$ starting at vertex 1 is covered.

We define the notion of weight formally as follows. Let $W(C, S)$ denote the walk induced by a traversal sequence $S \in \{0,1\}^*$ in a labeled chain $C \in \mathcal{L}(c)$ starting at vertex 1. A low reflection at vertex 0 represents a traversal from vertex $c$ to vertex 0 with $S$; and a high reflection at vertex $c$ is defined analogously. For a walk $W(C, S)$ terminating at vertex $v$, the ideal walk of $W(C, S)$, denoted by $\text{Ideal}(W(C, S))$, is the ideal walk in $C$ starting and terminating at vertices 1 and $v$, respectively, which has the same number of reflections as $W(C, S)$. The weight, $w(W(C, S))$, of a walk $W(C, S)$ is the length of the vertex-sequence of $\text{Ideal}(W(C, S))$.

The following two theorems justify the pruning of redundant branches in the quadtree.
Theorem 4. For a labeled chain $C \in \mathcal{L}(c)$ and two traversal sequences $S_1$ and $S_2$ of the same length-parity, if $wt(C, S_1) \leq wt(C, S_2)$, then for every $S \in \{0, 1\}^*$, $wt(C, S_1S) \leq wt(C, S_2S)$.

Proof. Omitted.

For $\alpha_1\alpha_2\cdots\alpha_n \in \{0, 1\}^n$, where $n \geq 0$, (1) $\pi_1(\alpha_1\alpha_2\alpha_3\alpha_4\cdots\alpha_n) = \alpha_1\overline{\alpha}_2\alpha_3\overline{\alpha}_4\cdots\alpha'_n$, where $\overline{\alpha}$ denotes the complement of $\alpha \in \{0, 1\}$, and $\alpha'_n$ is $\alpha_n$ if $n$ is odd, and $\overline{\alpha}_n$ otherwise, (2) $\pi_2(\alpha_1\alpha_2\alpha_3\alpha_4\cdots\alpha_n) = \alpha_1\alpha_2\overline{\alpha}_3\alpha_4\cdots\alpha''_n$, where $\alpha''_n = \overline{\alpha}_n$ if $n$ is odd, and $\alpha_n$ otherwise, and (3) $\pi_3 = \pi_2 \circ \pi_1$.

A more general form of redundancy is obtained via the functions $\pi_1$, $\pi_2$, and $\pi_3$. Let $1_{\{0, 1\}^*}$ denote the identity function on $\{0, 1\}^*$, and $\Pi = \{1_{\{0, 1\}^*}, \pi_1, \pi_2, \pi_3\}$.

Theorem 5. Let $C_i$ be a labeled chain in $\mathcal{L}(c)$, where $i \in \{0^{c-1}, 0^{c-2}1, \ldots, 1^{c-1}\}$, and $\pi \in \Pi$. For two traversal sequences $S_1$ and $S_2$ of even length-parity, if $wt(C_i, S_1) \leq wt(C_{\pi(i)}, S_2)$, then for every $S \in \{0, 1\}^*$, $wt(C_i, S_1S) \leq wt(C_{\pi(i)}, S_2\pi(S))$.

Proof. We modify the induction argument in the proof of Theorem 4 as follows. For the induction step, suppose that the statement in the theorem is true for every $S \in \{0, 1\}^*$ with $|S| \leq n$, where $n \geq 0$. Consider a sequence $S' \in \{0, 1\}^{n+1}$, and assume that $wt(C_i, S_1) \leq wt(C_{\pi(i)}, S_2)$. Consider the two cases for the parity of $n$.

The case of even $n$: We write $S' = Sa$, where $S \in \{0, 1\}^n$ and $a \in \{0, 1\}$. Applying the induction hypothesis on $S$, we have $wt(C_i, S_1S) \leq wt(C_{\pi(i)}, S_2\pi(S))$. Note that $\pi(S') = \pi(Sa) = \pi(S)\pi(a)$.

If $wt(C_i, S_1S) = wt(C_{\pi(i)}, S_2\pi(S))$, then the terminating vertices of the two walks $W(C_i, S_1S)$ and $W(C_{\pi(i)}, S_2\pi(S))$ have the same vertex-label — with odd parity due to the even length-parity of $S_1S$ and $S_2\pi(S)$. Starting at these two vertices with the same odd vertex-label in $C_i$ and $C_{\pi(i)}$, the sequence of vertices visited in $C_i$ induced by a traversal sequence $T \in \{0, 1\}^*$ is identical to the sequence of vertices in $C_{\pi(i)}$ induced by the sequence $\pi(T)$. Thus the traversal sequences $S_1Sa$ and $S_2\pi(S)\pi(a)$ in $C_i$ and $C_{\pi(i)}$, respectively, satisfy that $wt(C_i, S_1Sa) = wt(C_{\pi(i)}, S_2\pi(S)\pi(a))$.

If $wt(C_i, S_1S) < wt(C_{\pi(i)}, S_2\pi(S))$, then we have $wt(C_i, S_1S) + 2 \leq wt(C_{\pi(i)}, S_2\pi(S))$ as $S_1$ and $S_2$ are of the same length-parity. Notice that $|wt(C_i, S_1Sa) - wt(C_i, S_1S)| \leq 1$ and $|wt(C_{\pi(i)}, S_2\pi(S)\pi(a)) - wt(C_{\pi(i)}, S_2\pi(S))| \leq 1$, we have $wt(C_i, S_1Sa) \leq wt(C_i, S_1S) + 1 \leq wt(C_{\pi(i)}, S_2\pi(S)) - 1 \leq wt(C_{\pi(i)}, S_2\pi(S)\pi(a))$.

The case of odd $n$: We write $S' = Sa_1a_2$, where $S \in \{0, 1\}^{n-1}$ and $a_1, a_2 \in \{0, 1\}$. Applying the induction hypothesis on $S$, we have $wt(C_i, S_1S) \leq wt(C_{\pi(i)}, S_2\pi(S))$. Notice that $\pi(S') = \pi(Sa_1a_2) = \pi(S)\pi(a_1a_2)$. We proceed as in the previous case, and the sub-case when assuming that $wt(C_i, S_1S) = wt(C_{\pi(i)}, S_2\pi(S))$ is similar.
We now assume that \( wt(C_i, S_1 S) < wt(C_{\pi(i)}, S_2 \pi(S)) \). As the two sequences \( S_1 S \) and \( S_2 \pi(S) \) are of even length-parity, we have \( wt(C_i, S_1 S) + 2 \leq wt(C_{\pi(i)}, S_2 \pi(S)) \) (for each \( i \in \{0, 1\}^* \)).

Since \( |wt(C, Ta) - wt(C, T)| \leq 1 \) for all labeled chains \( C \in L(c), T \in \{0, 1\}^* \), and \( a \in \{0, 1\} \), we obtain that \( wt(C_i, S_1 Sa_1) \leq wt(C_{\pi(i)}, S_2 \pi(S) \pi(a_1)) \) and its strict inequality will give the desired result.

Consider now that \( wt(C_i, S_1 Sa_1) = wt(C_{\pi(i)}, S_2 \pi(S) \pi(a_1)) \). Both walks \( W(C_i), S_1 Sa_1 \) and \( W(C_{\pi(i)}, S_2 \pi(S) \pi(a_1)) \) terminate at vertices with the same even vertex-label \( v \) in \( C_i \) and \( C_{\pi(i)} \), respectively. If \( v \) is either 0 or \( c \), we obtain the desired result. Let \( i_v \) and \( \pi(i)_v \) denote the \( v \)th bits of the labels \( i \) and \( \pi(i) \) of \( C_i \) and \( C_{\pi(i)} \), respectively — the “forward” edge-labels associated with vertex \( v \) in \( C_i \) and \( C_{\pi(i)} \), respectively. Let \( a'_2 \) denote the last bit of \( \pi(a_1 a_2) \). We can verify that \( i_v = a_2 \) if and only if \( \pi(i)_v = a'_2 \). Thus both walks \( W(C_i, S_1 Sa_1 a_2) \) and \( W(C_{\pi(i)}, S_2 \pi(S) \pi(a_1 a'_2)) \) (\( = W(C_{\pi(i)}, S_2 \pi(S) \pi(a_1 a_2)) \)) terminate at vertices with the same vertex-label in \( C_i \) and \( C_{\pi(i)} \), respectively. Hence \( wt(C_i, S_1 Sa_1 a_2) = wt(C_{\pi(i)}, S_2 \pi(S) \pi(a_1 a_2)) \).

Combining the two cases for the parity of \( n \), we have \( wt(C_i, S_1 S') \leq wt(C_{\pi(i)}, S_2 \pi(S')) \), as desired. This completes the induction step. By induction, the theorem is proved.

For two traversal sequences \( S_1 \) and \( S_2 \) of even length-parity and \( |S_1| \geq |S_2| \), if for some \( \pi \in \Pi \), \( wt(C_i, S_1) \leq wt(C_{\pi(i)}, S_2) \) for each \( i \in \{0^{c-1}, 0^{c-2}1, \ldots, 1^{c-1}\} \), then by Theorem 6, for every \( S \in \{0, 1\}^* \), \( wt(C_i, S_1 S) \leq wt(C_{\pi(i)}, S_2 \pi(S)) \) for each \( i \in \{0^{c-1}, 0^{c-2}1, \ldots, 1^{c-1}\} \). This gives that \( \rho_1(S_1 S) \leq \rho_1(S_2 \pi(S)) \) for every \( S \in \{0, 1\}^* \). Thus \( S_1 \) is a redundant branch to explore.

We define a binary redundancy-ordering relation \( \leq_r \) on the set of even-length sequences of \( \{0, 1\}^* \) as follows. For \( S_1, S_2 \in \{0, 1\}^* \) of even length, \( S_1 \leq_r S_2 \) if and only if \( |S_1| \geq |S_2| \), and for some \( \pi \in \Pi \), \( wt(C_i, S_1) \leq wt(C_{\pi(i)}, S_2) \) for each \( i \in \{0^{c-1}, 0^{c-2}1, \ldots, 1^{c-1}\} \). Note that \( \leq_r \) is a preorder (that is, a reflexive and transitive binary relation), and every pair of sequences of length 2 are related under \( \leq_r \). We use the \( \leq_r \)-redundancy relation to prune the branches near the root of a quadtree, as well as locally at deeper depths.

### 3.2 The Shared Tree of Non-redundant Sequences

For a given chain length \( c \) and a pre-determined positive integer \( \Delta \), we generate a depth-\( \Delta \) tree \( T_\Delta \), consisting of all non-redundant pair-sequences (that is, pair-sequences maximal respect to \( \leq_r \)) of length at most 2\( \Delta \), inductively level by level. Suppose that we have constructed the tree \( T_\delta \) of all non-redundant pair-sequences of length at most 2\( \delta \), where \( 1 \leq \delta \leq \Delta \). Consider the set \( L_{\delta+1} \) of all immediate descendant pair-sequences of the leaves of \( T_\delta \). Denote by \( M(A) \) the set of all \( \leq_r \)-maximal elements in a set \( A \) of sequences of \( \{0, 1\}^* \). We compute the set \( M(V(T_\delta) \cup L_{\delta+1}) \) as follows:

1. Employ a simple divide-and-conquer algorithm to compute the set \( M(L_{\delta+1}) \), and
2. Compute the set $M(V(T_δ)∪L_{δ+1})$ by examining all pairs in $V(T_δ)×M(L_{δ+1})$ for $≤_r$-redundancy.

The non-redundant sequences in the shared tree $T_Δ$ are saved in a file and read each time the algorithm Build_Quadtree is using the chain length $c$. Before a pair-sequence of length at most $2Δ$ is examined, it is first looked up in the non-redundant sequence list; and if it is redundant, the sequence is not examined. At depth $δ$ that exceeds the pre-determined $Δ$, we perform small-scale local redundancy analysis to the subtrees rooted at depth $δ$ according to the $≤_r$-redundancy relation.

### 3.3 Local Trees of Non-redundant Sequences

For a given chain length $c$ and a pre-determined positive integer $Δ'$, and a pair-sequence $S$ of length $2(Δ+iΔ')$, where $i ≥ 0$, we can cover $S$ with the shared tree $T_Δ$ and a sequence/stack of local trees $t_1, t_2, \ldots, t_i$, each of which is a depth-$Δ'$ tree of locally non-redundant pair-sequences rooted at a prefix of $S$.

We can construct the sequence of local trees inductively on $i$ in a manner analogous to the construction of the shared tree $T_Δ$ described above. Before the local tree $t_{i+1}$ is generated, the pair-sequence at its root is compared to the non-redundant sequence lists of $t_i, t_{i-1}, \ldots, t_1$; and if it is redundant, it is treated as a leaf. (In our current implementation, only the non-redundant sequence list of $t_i$ is included for comparison.)

### 4 Conclusions

We have implemented the Quadtree Enumeration method with the improvements outlined in previous section (implementation is available upon request from the author), and have run the algorithm using a chain length of $c = 7$, a marks-to-pairs density upper bound of $ρ = \frac{64}{19}$, a $≤_r$-redundancy shared tree of depth $Δ = 10$, and local trees of depth $Δ' = 4$. This has resulted in $R(t, 7) ≥ 19t − 214$ and the length lower bounds for universal traversal sequences as in [8] with significant speedup, excluding the computation time of the $≤_r$-redundancy shared tree.

Our current effort is to employ the Quadtree Enumeration method incorporating the above-mentioned refinements and optimizations for a larger chain length of 9. Future work is directed at finding fast algorithms for solving the embedded Maximum Independent Set problems for induced subgraphs of inconsistency graphs, and at improving the construction of the $≤_r$-redundancy table.

### References


Competitive Facility Location along a Highway*

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Abstract. We consider a competitive facility location problem with two players. Players alternate placing points, one at a time, into the playing arena, until each of them has placed \( n \) points. The arena is then subdivided according to the nearest-neighbor rule, and the player whose points control the larger area wins. We present a winning strategy for the second player, where the arena is a circle or a line segment.

1 Introduction

The classical facility location problem [5] asks for the optimum location of a new facility (police station, super market, transmitter, etc.) with respect to a given set of customers. Typically, the function to be optimized is the maximum distance from customers to the facility — this results in the minimum enclosing disk problem studied by Megiddo [8], Welzl [12] and Aronov et al. [2].

Competitive facility location deals with the placement of sites by competing market players. Geometric arguments are combined with arguments from game theory to see how the behavior of these decision makers affect each other. Competitive location models have been studied in many different fields, such as spatial economics and industrial organization [1,9], mathematics [6] and operations research [3,7,11]. Comprehensive overviews of competitive facility locations models are the surveys by Friesz et al. [11], Eise1t and Laporte [3] and Eiselt et al. [4].

We consider a model where the behavior of the customers is deterministic in the sense that a facility can determine the set of customers more attracted to it than to any other facility. This set is called the market area of the facility. The collection of market areas forms a tessellation of the underlying space. If customers choose the facility on the basis of distance in some metric, the tessellation is the Voronoi Diagram of the set of facilities [10].

We address a competitive facility location problem that we call the Voronoi Game. It is played by two players, Blue and Red, who place a specified number, \( n \),

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of facilities in a region $U$. They alternate placing their facilities one at a time, with Blue going first. After all $2n$ facilities have been placed, their decisions are evaluated by considering the Voronoi diagram of the $2n$ points. The player whose facilities control the larger area wins.

More formally, let $\{b_i\}_{i=1}^n$ and $\{r_i\}_{i=1}^n$ be the respective locations of the blue and red points and set

$$B = |\{u \in U : \min_i d(u, b_i) < \min_i d(u, r_i)\}|,$$
$$R = |\{u \in U : \min_i d(u, r_i) < \min_i d(u, b_i)\}|$$

where $d(u, v)$ is an underlying metric and $|\cdot|$ indicates the area of a set. Blue wins if and only if $B > R$, Red wins if and only if $R > B$ and the game ends in a tie if $B = R$.

The most natural Voronoi Game is played in a two-dimensional arena $U$ using the Euclidean metric. Unfortunately nobody knows how to win this game, even for very restricted regions $U$. In this note we present strategies for winning one-dimensional versions of the game, where the arena is a circle or a line segment, and variations. In other words, we consider competitive facility location along an Australian highway.

The next section discusses the simplest game, on the circle. It is obvious that the second player, Red, can always achieve a tie by playing on the antipode of Blue’s move. One might try to tweak this strategy such that it results in a win for Red. This doesn’t seem to work, and we present instead a quite different winning strategy for Red.

Section 3 describes how this strategy remains a winning strategy even if the rules of the game are drastically relaxed.

In Section 4 we finally turn to the line segment arena. It would appear that Blue has an advantage here, because it can play the midpoint of the segment in its first move. We show that this doesn’t help, and prove that Red still has a winning strategy. The strategy is quite similar to the one for the circle case, but its analysis (because of a loss of symmetry) is more detailed.

## 2 The Basic Circle Game

There are two players, Blue and Red, each having $n$ points to play, where $n > 1$. They alternate placing these points on circle $C$, with Blue placing the first point, Red the second, Blue the third, etc., until all $2n$ points are played. We assume that points cannot lie upon each other. Let $\{b_i\}_{i=1}^n$ be the locations of the blue points and $\{r_i\}_{i=1}^n$ be those of the red ones. After all of the $2n$ points have been played each player receives a score equal to the total circumference of the circle that is closer to that player than to the other, i.e., Blue and Red have respective scores

$$B = |\{x \in C : \min_i d(x, b_i) < \min_i d(x, r_i)\}|,$$
$$R = |\{x \in C : \min_i d(x, r_i) < \min_i d(x, b_i)\}|$$
The player with the highest score (the larger circumference) wins.

The question that we address here is, *Does either player have a winning strategy and, if yes, what is it?* We will see below that the second player, Red, always has a winning strategy.

Before giving the strategy we introduce some definitions. We parameterize the circle using the interval $[0, 1]$, where the points 0 and 1 are identified. Arcs on the circle are written as $[x, y]$ implying the clockwise arc running from $x$ to $y$, as in $[.5, .6]$ or $[.9, .1]$.

**Definition 1.** The $n$ points $u_i = \frac{i}{n}$, $i = 0, 1, \ldots, n - 1$ are keypoints.

We call an arc between two clockwise consecutive red/blue points an interval. The interior of an interval is free of red/blue points. At any given time during the game the circle is partitioned into intervals. An interval is monochromatic if its endpoints have the same color, and bichromatic if they have different colors. A blue interval is a blue monochromatic one, a red interval a red monochromatic one. We denote the total length of all red intervals by $R_m$, and the total length of all the blue intervals by $B_m$. An interval is called a key interval if both of its endpoints are keypoints.

The important thing to notice is that at the end of the game the length of each bichromatic interval is divided equally among the two players, so $R - B = R_m - B_m$ and Red wins if and only if $R_m > B_m$. We devise our strategy to force this to happen.

Since we can parameterize the circle arbitrarily, we can assume without loss of generality that Blue plays his first point on 0 and thus on a keypoint. We now describe Red’s winning strategy. Figure 1 shows an example.

**Red’s Keypoint Strategy**

**Stage I:** If there is an empty keypoint then Red plays onto the keypoint. Stage I ends after the last keypoint is played (by either Red or Blue).

**Stage II:** If there is no empty keypoint and it is not Red’s last move then Red plays her point into a largest blue interval. We call this breaking the blue interval. Stage II ends when Blue plays his last point.

**Stage III:** Red’s last move. There are two possibilities:

(i) if there exists more than one blue interval then Red breaks a largest one by placing her point inside.

(ii) if there is only one blue interval define $\ell < \frac{1}{n}$ to be its length. Red’s move is to go to a bichromatic key interval and claim a red interval of length larger than $\ell$ by placing a red point closer than $\frac{1}{n} - \ell$ to the blue endpoint of the bichromatic key interval.

The two following lemmas will be needed.

**Lemma 1.** Let $B$ be a set of $b$ blue points and let $R$ be a set of $r$ red points currently on the circle with $b \geq r$. Let $n(R)$ be the number of red intervals they form and $n(B)$ the number of blue ones. Then $n(B) - n(R) = b - r$. 
Fig. 1. There are four points to be played for both Blue and Red. The white dots represent Blue’s points and the black dots represent Red’s points. We label the dots in chronological order.

Proof. The proof will be by induction on \( r \). If \( r = 0 \) then \( b \) blue points form \( b \) blue intervals so \( n(B) = b \), \( n(R) = 0 \) and the condition \( n(B) - n(R) = b - r \) is satisfied.

Now suppose that the lemma is true for all configurations of \( b \) blue points and \( r - 1 \) red ones. Deleting any red point \( p \) from \( R \) leaves \( b \) blue and \( r - 1 \) red points, so \( n(B) - n(R - \{ p \}) = b - r + 1 \). We now add the red point \( p \) back into the configuration and ask how the monochromatic intervals can change. There are three possible placements of \( p \):

- (i) inside a red interval, increasing \( n(R) \) by one and leaving \( n(B) \) unchanged.
- (ii) inside a blue interval, decreasing \( n(B) \) by one and leaving \( n(R) \) unchanged.
- (iii) inside a bichromatic interval, increasing \( n(R) \) by one and leaving \( n(B) \) unchanged.

After all three of these cases we find that \( n(B) - n(R) = n(B) - n(R - \{ p \}) - 1 = b - r \).

Lemma 2. Suppose that all \( n \) keypoints are covered and Blue has just moved (possibly covering the last keypoint). If there is only one blue interval and this interval has length \( < \frac{1}{n} \), then there exists a bichromatic key interval.

Proof. We apply the pigeon hole principle: At most \( 2n - 1 \) points have been played, \( n \) of them on keypoints. Consider the \( n \) circle arcs of length \( \frac{1}{n} \) formed by the \( n \) keypoints. Since the blue interval has length \( < \frac{1}{n} \), at least one of its endpoints is inside an arc. That leaves only \( n - 2 \) points to have been played inside the \( n - 1 \) remaining arcs. Therefore, one of the arcs must be free of points, forming a key interval. Since there is only one blue interval, there is no red interval by Lemma 1. Therefore, this key interval is bichromatic.

Theorem 1. The keypoint strategy is a well-defined winning strategy for Red.

Proof. We start with a simple observation. Since the circle contains only \( n \) keypoints and Blue’s first move covers the first keypoint, Red will play onto at most \( n - 1 \) keypoints. Thus Stage I always ends before Red plays her last point.
Consider Stage II. Lemma 1 implies that after each play by Blue \((b = r + 1)\) there is always at least one blue interval on the circle, so Stage II of the strategy is indeed well defined.

We make two observations concerning the situation after Stage II, when Red has played her \(n - 1\)’st point. The first is that there is no blue key interval. Let \(k\) be the number of keypoints played by Blue during the game. Red has covered the remaining \(n - k\) keypoints by the end of Stage I. If \(k = 1\) (the only case in which Red skips Stage II), then there certainly is no blue key interval as there is only one blue keypoint. When \(k > 1\), Blue can define at most \(k - 1\) blue key intervals with its \(k\) keypoints (since Red has at least one keypoint). Note that since all keypoints are played by the end of Stage I, all intervals in Stages II and III have length at most \(\frac{1}{n}\). In particular a blue key interval is longer than any other blue interval. Since Red plays \(k - 1\) points in Stage II, all blue key intervals are broken during Stage II.

The second observation concerning the situation after stage II is that all red intervals are key intervals. This statement is true at the end of Stage I, as Red has so far only played onto keypoints, and all keypoints are covered. During Stage II, Red uses her points to break blue intervals, and therefore creates bichromatic intervals only. Blue cannot create red intervals, and so, at the end of Stage II, all red intervals are indeed key intervals.

We now show why Stage III is well defined and why Red wins. Suppose that Blue has just played his last point and it is now time for Stage III, Red’s last move. From Lemma 1 we know that \(n(B) \geq 1\).

If \(n(B) > 1\) before Red’s last move then the strategy is well defined: Red breaks a largest blue interval. This decreases \(n(B)\) by 1 so the game ends with \(n(B) \geq 1.\) By Lemma 1 we have \(n(R) = n(B) \geq 1.\) But now note that from the observations in the preceding paragraphs all existing red intervals are key intervals while all existing blue intervals have length strictly less than \(\frac{1}{n}\). Since all red intervals are longer than all blue intervals and there are the same number of red ones and blue ones we find that \(R_m > B_m\) and Red wins.

If \(n(B) = 1\) before Red’s last move the strategy requires that the unique blue interval has length \(\ell < \frac{1}{n}\), and that there exists a bichromatic key interval. The first fact was already observed above, the second fact follows from Lemma 2.

After Red places her last point Blue still has one blue interval of length \(\ell\) while Red has one red interval of length \(> \ell\). Thus \(R_m > \ell = B_m\) and Red wins.

### 3 A Modified Circle Game

The basic game can be modified in many different ways. The simplest modification allows the players to play more than one point at a time. More complicated modifications permit the players (both or one) to choose before each turn, how many points they play.

Suppose that there are \(k \leq n\) rounds. Let \(\beta_i\) and \(\gamma_i\) be the numbers of points that Blue and Red play respectively in round \(i\). Suppose that the following restrictions are placed.
– $\forall 1 \leq i \leq k, \beta_i, \gamma_i > 0$.
– $\forall 1 \leq j \leq k, \sum_{i=1}^{j} \beta_i \geq \sum_{i=1}^{j} \gamma_i$.
– $\sum_{i=1}^{k} \beta_i = \sum_{i=1}^{k} \gamma_i = n$.
– $\beta_1 < n$.

Then Red still wins by following exactly the same strategy as in the previous section of first filling in the keypoints and then breaking the largest blue intervals until Red plays it’s last point when it follows the Stage III rules. The proof that the strategy is well defined and wins is almost exactly the same as the one in the previous section so we will not repeat it here.

Note that this generalization includes both the original game and the “batched” version in which each player plays the same number ($> 1$) of points at each turn. Note, too, that $k$, $\beta_i$, and $\gamma_i$ need not be fixed in advance. For example, Blue may decide at every move how many points he will play and then Red plays the same number.

We conclude this section by noting that the condition $\beta_1 < n$ is essential since otherwise Blue would play the keypoints, forcing a tie.

4 The Line Segment Version

We now move on to the version of the game played on a line segment. We consider it to be horizontal and parameterized as $[0, 1]$. The scoring is the same as in the basic circle game except that the player with the leftmost point claims everything between 0 and the point, and the player with the rightmost point claims everything between the point and 1. We assume that $n > 1$, and points cannot lie upon each other. When $n = 1$, Blue wins by placing onto $\frac{1}{2}$.

We modify some of the old definitions and introduce new ones:

**Definition 2.** The $n$ points $u_i = \frac{i}{2n} + \frac{i}{n}$, $i = 0, 1, \ldots, n - 1$ are keypoints.

The left segment is the segment from 0 to the leftmost red or blue point. The right segment is the segment from the rightmost red or blue point to 1. The border interval is the union of the left and right segments. An interval is a section of the line segment with red/blue endpoints and no red/blue points in its interior. We consider the border interval an interval. An interval, including the border interval, is monochromatic if its endpoints have the same color, and bichromatic if they have different colors. With this definition of intervals, Lemmas 11 and 2 are true for the line segment as well.

We denote the total length of all of the blue intervals—including, if appropriate, the border interval—by $B_m$, the total length of all of the red intervals—again including, if appropriate, the border interval—by $R_m$. When the border interval is bichromatic, we use $B_b$ to denote the length of the left/right segment with a blue endpoint and $R_b$ to denote the length of the left/right segment with a red endpoint. If the border interval is monochromatic, then $B_b = R_b = 0$. Since all bichromatic non-border intervals are equally shared by both players $R - B = (R_m + R_b) - (B_m + B_b)$ and, as in Section 2, we design our strategy so that Red finishes with the right hand side of the equation $> 0$. 
We now introduce the line strategy, a modified version of the circle strategy. Figure 2 shows an example.

Red’s Line strategy

Stage I: If there is an empty keypoint then Red plays the keypoint. If \(u_0\) or \(u_{n-1}\) have not yet been played then Red should play onto one of them first. Stage I ends after the last keypoint is played by either Red or Blue. Note that the game may finish in Stage I.

Stage II: If there is no empty keypoint and it is not Red’s last move then

(i) if there exists at least one blue non-border interval, then Red should break a largest blue non-border interval by placing her point inside.

(ii) if the border interval is the only blue interval, then there are two possible cases:
   (a) One of the blue endpoints of the blue border interval is a keypoint:
   Without loss of generality assume that it is \(u_0\) (the other case is symmetric) and the other endpoint is \(1 - \ell\). From Stage I the other endpoint cannot be the keypoint \(u_{n-1}\) so \(\ell < \frac{1}{2n}\). Red now places her new point at \(x\) where \(x\) is anywhere in \((\ell, u_0]\).
   (b) Neither of the endpoints of the border interval are keypoints:
   Let \(\ell\) be the length of the blue border interval; \(\ell < \frac{1}{n}\). There must then exist a bichromatic key interval (Lemma 2). Red places her new point in that interval to form a new red interval of length > \(\ell\).

Stage III: If Red is placing her last point, we have two mutually exclusive cases:

(i) if there exists more than one blue interval, then Red should break a largest non-border one.

(ii) if there exists only one blue interval, then let its length be \(\ell\); we will see below that \(\ell < \frac{1}{n}\). Red should go to a bichromatic key interval (one will exist from Lemma 2) and claim a red interval of length > \(\ell\) as follows.
   • If the bichromatic key interval is not the border one, Red can do this by creating a new red interval of length > \(\ell\).
   • If the bichromatic key interval is the border one, then Red already possesses \(\frac{1}{2n}\) of it because it has all of either \([0, u_0]\) or \([u_{n-1}, 1]\). Red can therefore go to the other side it does not possess, and grab enough length to have a red border interval of length > \(\ell\).

Theorem 2. The line strategy is a well-defined winning strategy for Red.
Proof. Note that this strategy differs in at least one major aspect from the circle strategy: since we have lost circular symmetry it cannot be guaranteed that Blue plays onto at least one keypoint, and so it is possible that the game will end in Stage I, with Red playing all $n$ keypoints. In this case, all red intervals (including, possibly, the border interval) are key intervals and all blue intervals have length $< \frac{1}{n}$. By Lemma 1 Blue and Red have the same number of monochromatic intervals, so $B_m < R_m$. If the border interval is monochromatic, then $B_b = R_b = 0$ and Red wins. If the border interval is bichromatic, then one of its endpoints must be the red point $u_0$ or $u_{n-1}$. This implies that $B_b < R_b = \frac{1}{2n}$, and Red wins.

In what follows we may therefore assume that Blue plays onto at least one keypoint during the game. We will show that at the end of the game it will always be true that $B_m < R_m$ and, if the border interval is bichromatic, then $B_b \leq R_b$. The theorem will follow.

First note that under this assumption Stage I always ends with all keypoints covered, and Stage III is reached. Note further that Red’s first move is onto either $u_0$ or $u_{n-1}$.

We consider Stage II. After Blue’s every move there exists at least one blue interval (possibly the border interval) by Lemma 1. If there is only one such blue interval, there is no red interval. Thus, one of the two conditions (i) or (ii) of Stage II holds. The strategy is clearly well defined in cases (i) and (ii)(a), the validity of (ii)(b) follows from Lemma 2.

We will need one more observation.

Lemma 3. After Blue’s last move, there is no blue key interval.

Proof. Let $k$ be the number of keypoints played by Blue, where $1 \leq k \leq \lceil \frac{n}{2} \rceil$. There are therefore at most $k - 1$ blue key intervals after Stage I. Red occupies $n - k$ keypoints in Stage I, and so Stage II lasts for $k - 1$ rounds. This is sufficient for all blue key intervals to be broken, since a blue key interval is longer than any other blue interval in Stage II (i) and Stage II (ii)(a).

We now prove that Red wins. There are two cases.

Lemma 4. Assume that case (ii) of Stage II never occurs. Then all red intervals are key intervals after Blue’s last move.

Proof. After Stage I all red intervals are key intervals, since Red has only played keypoints and all keypoints are covered. During Stage II
Red uses all her points to break blue intervals (since case (ii)(b) does not occur), and so creates only bichromatic intervals. As Blue cannot create red intervals, all red intervals remaining after Stage II are indeed key intervals.

We examine the result of Stage III, assuming that case (ii) of Stage II did not occur. Assume first that Red plays case (i) of Stage III. There are then equal numbers of red and blue intervals left after the last move. Since all blue intervals have length < \( \frac{1}{n} \) by Lemma 3 and all red intervals are key intervals by Lemma 4, we have \( B_m < R_m \). If the border interval is monochromatic, then \( B_b = R_b = 0 \) and Red wins. If the border interval is bichromatic, then its red endpoint must be \( u_0 \) or \( u_{n-1} \) (since Red plays case (i) of Stage II only), and so \( B_b \leq R_b = \frac{1}{2n} \) and Red wins.

On the other hand, assume now that Red plays case (ii) of Stage III. Then Blue has total length \( \ell \), Red has > \( \ell \), and so Red wins.

We now consider the remaining case, where case (ii) of Stage II does occur.

**Lemma 5.** Assume that case (ii) of Stage II occurs at least once. After Red’s last move in Stage II, there is no blue non-border interval and \( B < R \).

**Proof.** We prove that the statement is true after the last occurrence of case (ii) in Stage II, and after each subsequent move by Red.

Consider the last occurrence of case (ii). Before Red’s move, the border interval is the only blue interval.

If Red plays case (ii)(a), there is no blue interval at all after Red’s move. We have then \( B_m = R_m = 0 \) and \( B_b < R_b \), which implies \( B < R \).

If Red plays case (ii)(b), Red claims a red interval longer than the blue border interval. So after Red’s move we have \( B < R \), as \( B_m < R_m \) and \( B_b = R_b = 0 \).

Consider now the remaining moves of Stage II. In all these moves Red plays case (i), and so we can deduce that Blue uses his move to create a new blue non-border interval. Red immediately breaks this blue interval. This leaves \( B_m, R_m, B_b, \) and \( R_b \) unchanged and destroys the only blue non-border interval. The claim therefore remains true after each subsequent move by Red, and in particular after Red’s last move in Stage II.

We consider the situation right before Blue’s last move. By Lemma 5, there is no blue non-border interval, and \( B < R \).

If the border interval is blue before Blue’s last move and Blue uses that move to create a blue non-border interval move, Red plays case (i) of Stage III. This breaks the new blue non-border interval, returning \( B \) and \( R \) to their state before Blue’s last move, and so Red wins.

In all other cases there is a single blue interval after Blue’s last move. It cannot be a key interval by Lemma 3 and so, by Lemma 2, there is a bichromatic key interval. Red claims a red interval longer than the blue interval, and wins.
5 Conclusions

We have given strategies for one-dimensional competitive facility location, allowing the second player, Red, to win. The margin by which Red wins is very small, however, and in fact Blue can make it as small as he wants. Moreover, our arguments can easily be transformed into a strategy that allows the first player, Blue, to lose by an arbitrarily small margin. Do our findings have any bearing on the two-dimensional Voronoi Game? The concept of key points turned out to be essential to our strategies. Is there a similar concept in two dimensions? A natural attempt would be to define a uniform square grid of key points. Perhaps surprisingly, a player governing this grid can still lose the game by a considerable margin.

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References

Membership for Core of LP Games and Other Games*

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Abstract. Let \( \Gamma \equiv (N, v) \) be a cooperative game with the player set \( N \) and characteristic function \( v : 2^N \rightarrow R \). An imputation of the game is in the core if no subset of players could gain advantage by splitting from the grand coalition of all players. It is well known that, for the linear production game, and the flow game, the core is always non-empty (and a solution in the core can be found in polynomial time). In this paper, we show that, given an imputation \( x \), it is \( NP \)-complete to decide it is not a member of the core, in both games. The same also holds for Steiner tree game. In addition, for Steiner tree games, we prove that testing the total balancedness is \( NP \)-hard.

Key words: cooperative game, core, network flow, linear programming, Steiner tree, \( NP \)-completeness.

1 Introduction

A cooperative game \( \Gamma = (N, v) \) consists of a player set \( N = \{1, 2, \cdots, n\} \) and a characteristic function \( v : 2^N \rightarrow R \). For a revenue game \( \Gamma \), the core is defined by

\[
\text{Core}(\Gamma) = \{ x \in R^n : x(N) = v(N) \text{ and } x(S) \geq v(S), \forall S \subseteq N \},
\]

where \( x(S) = \sum_{i \in S} x_i \). The study of the core are closely associated with another important concept, the balanced set. The collection \( \mathcal{B} \) of subsets \( S \) of \( N \) is balanced if there exists a set of positive numbers \( \gamma_S, S \in \mathcal{B} \), such that for each \( i \in N \), we have \( \sum_{S \in \mathcal{B}} \gamma_S x_i = 1 \). A game \( (N, v) \) is called balanced if \( \sum_{S \in \mathcal{B}} \gamma_S v(S) \leq v(N) \) holds for every balanced collection \( \mathcal{B} \) with weights \( \{ \gamma_S : S \in \mathcal{B} \} \). With techniques essentially the same as the linear programming duality, Shapley [13] proved that a game has a non-empty core if and only if it is balanced. For a subset \( S \subseteq N \), we define the induced subgame \( (S, v_S) \) on \( S : v_S(T) = v(T) \) for

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every subset $T$ of $S$. A game is called totally balanced if all its subgames are balanced, i.e., all its subgames have non-empty cores.

The computational complexity as a rationality measure for game theoretical concepts has attracted more and more attention recently. In cooperative game studies, work has focused on games that have a succinct representation, that is, the game value is defined as a solution to a combinatorial optimization problem, for which the input size is polynomial in the number of players, since first suggested by Megiddo \[11\]. Various interesting complexity structures have started to emerge as a result, especially for the study of core, which will be the focus of our paper. Deng and Papadimitriou \[4\] found a game for which the core is nonempty if and only if a certain imputation (Shapley value in this case) is in the core. Faigle, et al. \[5\] show that, for the minimum cost spanning tree (MCST) game, membership testing is $NP$-complete. It is known, however, the MCST game has a non-empty core and a member in the core can be found in polynomial time \[8\]. Goemans and Skutella \[7\] recently showed that, for a facility location game, if the core is non-empty, a solution in core can be found in polynomial time, and membership testing can be done in polynomial time. However, it is $NP$-complete to decide if a core is not empty. In this paper, we consider the flow game and the linear production game. Both are known to have a non-empty core and a solution in the core can be found in polynomial time \[9,10,12,2\]. However, we show that, given a solution, it is $NP$-complete to tell it is not in the core.

The flow game was introduced by Kalai and Zemel \[9,10\]. It is associated with the maximum flow from a source to a sink on a network, where each player controls one arc in the network. The value of a subset of players is the same as the maximum flow from the source to the sink on the subgraph induced by the edges controlled by that subset of players. The flow game is totally balanced, and on the other hand, every non-negative totally balanced game can be formulated as a flow game. Owen \[12\] introduced a linear production game in which each player $j$ controls a certain resource vector $b_j$. Jointly, they maximize a linear objective function $cx$, subject to the resource constraint $Ax \leq \sum_{j \in S} b_j$. The value a subset $S$ of players can achieve on their own is the maximum they can achieve with resources collectively owned by this subset: $\max \{ cx : Ax \leq \sum_{j \in S} b_j \}$. Owen showed that the core for linear production game is always nonempty and one imputation in the core can be immediately obtained from his proof using variables of the dual linear program. Curiel \[1\] proved that the class of linear programming games is also equivalent to the class of totally balanced games. Therefore, the class of linear programming games is equivalent to the class of flow games. These reductions for the equivalence proof, however, involve in exponential time and space in the number of players. Therefore, computational complexity results for these two games may not be equivalent.

In Section 2, we first show that membership test in the core is $co-NP$-complete for flow games. Because of the special structure we have for the reduction, the result flow game has a polynomial size linear production game formulation. Therefore, the same conclusion for linear production games follows as a result. In Sec-
tion 3, we consider the same problem for the Steiner tree game and show it is also NP-hard.

In Section 4, we show it is NP-hard to decide whether a Steiner tree game is total balanced. Naturally, linear production game and flow game are total balanced. Deng, et al., considered total balanceness of several interesting combinatorial optimization games [3]. Most were shown to be polynomially decidable. The worst complexity is for a coloring game, for which total balancedness is equivalent to a graph being perfect graph. Our result is the first example of NP-hardness for total balanced condition.

2 Complexity of Membership Test in the Core for Flow Games and Linear Production Games

Theorem 2.1 It is NP-complete to show that, given a maximum flow game $\Gamma_D = (E, v)$ on network $D = (V, E; \omega)$ and a vector $x \in R^n$ ($n = |E|$) with $x(E) = v(E)$, is $x$ not a core member of the game, i.e., does there exist a coalition $S \subseteq E$ such that $x(S) < v(S)$?

Proof. The problem is in $NP$ because a coalition $S \subseteq E$ with $x(S) < v(S)$ (if it exists) can be exhibited and checked easily. To prove completeness, we establish a polynomial transformation from a basic NP-complete problem, EXACT COVER BY 3-SETS [6]:

Given a finite set $X = \{x_1, x_2, \ldots, x_{3q}\}$ and a collection $F = \{f_1, f_2, \ldots, f_{|F|}\}$ of 3-element subsets of $X$ ($|F| \geq q$), does $F$ contain an exact cover for $X$, that is, does there exist a subcollection $F' \subseteq F$ such that every element of $X$ occurs in exactly one member of $F'$?

For an instance of X3C, we construct a maximum flow game $\Gamma_D$ on network $D = (V, E; \omega)$:

Thus, the vertex set of $D$ is $V = V_X \cup V_F \cup \{s, t, v_0, v'_0\}$. (See Fig.1).

- $V_X = \{v_1, v_2, \ldots, v_{3q}\}$, one element-vertex $v_i \in V_X$ for each $x_i \in X$, $i = 1, 2, \ldots, 3q$;
- $V_F = \{u_1, u_2, \ldots, u_{|F|}\}$, one set-vertex $u_j \in V_F$ for each $f_j \in F$, $j = 1, 2, \ldots, |F|$;
- source $s$, sink $t$, two additional vertices $v_0$ and $v'_0$.

The arc set $E$ and the corresponding capacity function $\omega : E \to R^+$ are given as follows:

- $a_0 = (s, v_0)$, $\omega(a_0) = 6q|F|$;
- $b_0 = (v'_0, t)$, $\omega(b_0) = 3|F| - 3q$;
- $A_1 = \{(v_0, u_i) : i = 1, 2, \ldots, |F|\}$, $\forall e \in A_1 : \omega(e) = 3$;
- $A_2 = \{(u_i, v_j) : x_j \in f_i, i = 1, 2, \ldots, |F|, j = 1, 2, \ldots, 3q\}$, $\forall e \in A_2 : \omega(e) = 1$;
- $A_3 = \{(v_j, t) : j = 1, 2, \ldots, 3q\}$, $\forall e \in A_3 : \omega(e) = 1$;
- $B = \{(u_i, v'_0) : i = 1, 2, \ldots, |F|\}$, $\forall e \in B : \omega(e) = 3$. 
The arc set of $D$ is $E = \{a_0, b_0\} \cup A_1 \cup A_2 \cup A_3 \cup B$, $n = |E|$, (See Fig. 1) Consider maximum flow game $\Gamma_D$ on this network $D$. Obviously, $v(E) = 3|F|$.

Let $x \in R^n$ be a candidate for a core member, which is given by:

$$
\begin{align*}
x(a_0) & = 2q - \frac{1}{3|F|}; \\
x(b_0) & = 3|F| - 3q + \left(\frac{1}{3|F|} - \frac{1}{3(q + 1)}\right); \\
x(e) & = \frac{1}{3|F|(q + 1)}, \quad \forall e \in A_1; \\
x(e) & = 0, \quad \forall e \in A_2; \\
x(e) & = \frac{1}{3}, \quad \forall e \in A_3; \\
x(e) & = 0, \quad \forall e \in B.
\end{align*}
$$

Easily, we have $x(E) = \sum_{e \in E} x(e) = 3|F| = v(E)$. We shall claim that $x$ is not in the core of the maximum flow game $\Gamma_D$ if and only if $F$ contains an exact cover for $X$.

Suppose that $x$ is not in the core of $\Gamma_D$, and let $S \subset E$ be a coalition such that $x(S) < v(S)$. We would like such an ‘unsatisfied’ coalition $S$ to exhibit an exact cover.

1). $S$ contains $a_0$. Otherwise, $x(S) \geq 0 = v(S)$, a contradiction.
2). $S$ does not contain $b_0$. Otherwise, if $b_0 \in S$, we will discuss in two cases: Case 1. $|S \cap A_3| = k \leq 3q - 1$.

$$
v(S) \leq k + (3|F| - 3q),$$

$$
x(S) \geq (2q - \frac{1}{3|F|}) + [3|F| - 3q + \left(\frac{1}{3|F|} - \frac{1}{3(q + 1)}\right)] + \frac{1}{3}k > v(S).
$$

Case 2. $|S \cap A_3| = 3q$. 

Fig. 1. Network $D = (V, E; \omega)$. 
If $|S \cap A_1| \leq |F| - 1$, then $v(S) \leq 3|F| - 3$,

$$x(S) \geq (2q - \frac{1}{3|F|}) + |F| - 3q + \left(\frac{1}{3|F|} - \frac{1}{3(q + 1)}\right) + \frac{1}{3} > v(S).$$

If $|S \cap A_1| = |F|$, then $v(S) \leq 3|F|$,

$$x(S) = (2q - \frac{1}{3|F|}) + |F| - 3q + \left(\frac{1}{3|F|} - \frac{1}{3(q + 1)}\right) + \frac{1}{3} \geq v(S).$$

It is contrary to the fact $x(S) < v(S)$ in both Case 1 and Case 2, so $b_0 \not\in S$.

3). $S$ contains all $3q$ arcs in $A_3$, and $v(S) = 3q$.

If $|S \cap A_3| = k \leq 3q - 1$, then

$$x(S) \geq (2q - \frac{1}{3|F|}) + \frac{1}{3} k \geq k + \left(\frac{2}{3} - \frac{1}{3|F|}\right) > k \geq v(S),$$

which is a contradiction. So $|S \cap A_3| = 3q$.

If $v(S) < 3q$, then we have $v(S) \leq 3q - 1$, since $v(S)$ must be integer under the given integral arc capacities. It follows that

$$x(S) \geq (2q - \frac{1}{3|F|}) + 3q \frac{1}{3} > 3q - 1 \geq v(S),$$

also a contradiction. So $v(S) = 3q$.

4). $S$ contains exactly $q$ arcs in $A_1$. Suppose that $|S \cap A_1| = k$, then

$$x(S) = (2q - \frac{1}{3|F|}) + k - \frac{1}{3|F|} + 3q \frac{1}{3} = 3q - \frac{1}{3|F|} \left(1 - \frac{k}{q + 1}\right).$$

Since $x(S) < v(S) = 3q$, we have $k < q + 1$. Again since each arc in $A_1$ has only 3 arcs in $A_2$ followed, so we must have $k \geq q$. That is, $k = q$.

Therefore, under our assumption of $x \not\in \text{Core}(I_D)$, the above discussions 1)—4) imply that $F$ contains an exact cover of $X$.

On the other hand, if $F$ admits an exact 3-cover $F' = \{f_{i_1}, f_{i_2}, \cdots, f_{i_q}\}$, let $S = \{a_0\} \cup \{(v_0, u_{i_k}) : k = 1, 2, \cdots, q\} \cup A_2 \cup A_3$, then

$$v(S) = 3q,$$

$$x(S) = (2q - \frac{1}{3|F|}) + \frac{1}{3|F|} + q \frac{1}{3|F|} > q < 3q,$$

which implies that $x$ is not in the core of the maximum flow game $I_D$.

Finally, easily to see that the construction of $I_D$ and the candidate core member $x$ can obviously be carried out in polynomial time. The proof is completed.

As a direct corollary, we come to the same conclusion for linear production games in Corollary 2.2, that is, the problem of checking membership in the core is also co-NP-complete.

**Corollary 2.2** It is NP-complete to decide that given a linear production game $I \equiv (N, v)$ and a vector $x \in R^n \ (n = |N|)$ with $x(N) = v(N)$, is $x$ not a core member of this game, i.e., does there exist a coalition $S \subset N$ such that $x(S) < v(S)$?
3 Checking Membership of Core for Steiner Tree Game

**Theorem 3.1** It is NP-hard to decide that given a balanced Steiner tree game $\Gamma(G) = (N, v)$ on weighted graph $G = (V, E; \omega)$ and a vector $x \in \mathbb{R}^n$ ($n = |N|$) with $x(N) = v(N)$, does there exist a coalition $S \subseteq N$ such that $x(S) > v(S)$.

Proof. Given an arbitrary instance of EXACT COVER BY 3-SETS (X3C): a finite set $X = \{x_1, x_2, \ldots, x_n\}$ and a collection $F = \{f_1, f_2, \ldots, f_{|F|}\}$ of 3-element subsets of $X$ ($q \geq 2, |F| \geq 2q$). We construct a Steiner tree game $\Gamma(G)$ on weighted graph $G = (V, E; \omega)$ and a candidate vector $x$ for a core member.

The vertex set $V$ of $G$ consists of four parts:

- $N_X = \{v_1, v_2, \ldots, v_{3q}\}$, the vertex $v_i \in N_X$ is corresponding to the element $x_i \in X$, $i = 1, 2, \ldots, 3q$, $N_X$ represents a subset of consumers;
- $N_F = \{u_1, u_2, \ldots, u_{|F|}\}$, the vertex $u_j \in N_F$ is corresponding to the subset $f_j \in F$, $j = 1, 2, \ldots, |F|$, $N_F$ also represents a subset of consumers;
- $g_1$ and $g_2$ are two additional vertices, they represent switches;
- $v_0$ represents central supplier.

Thus, $V = N_X \cup N_F \cup \{g_1, g_2, v_0\}$. Let $N = N_X \cup N_F$, $n = |N|$. (See Fig.2).

The edge set $E$ of $G$ and edge weight function $\omega : E \rightarrow \mathbb{R}^+$ are defined as follows:

- $a_0 = (v_0, g_1)$, $\omega(a_0) = 2q - 1$;
- $b_0 = (g_1, g_2)$, $\omega(b_0) = q + 1$;
- $E_1 = \{(g_1, u_i) : i = 1, 2, \ldots, |F|\}$, $\forall e \in E_1 : \omega(e) = q + 1$;
- $E_2 = \{(g_2, u_i) : i = 1, 2, \ldots, |F|\}$, $\forall e \in E_2 : \omega(e) = q$;
- $E_3 = \{(u_i, v_j) : x_j \in f_i, i = 1, 2, \ldots, |F|, j = 1, 2, \ldots, 3q\}$, $\forall e \in E_3 : \omega(e) = 2q + 1$;

The edge set $E = \{a_0, b_0\} \cup E_1 \cup E_2 \cup E_3$. (See Fig.2). Denote the Steiner tree game on the weighted graph $G$ by $\Gamma(G) = (N, v)$.

Our construction of graph $G$ is much similar to that in Faigle, et al.[5]. The only difference is the weights of edges in $E_3$. The weights of edges in $E_3$ is sufficient large so that each vertex in $N_X$ does not serves as a switch in the minimum Steiner tree of $G$. For any coalition $S \subseteq N$, denote the minimum Steiner tree w.r.t. $S \cup \{v_0\}$ by $T_S = (V_S, E_S)$. We have

1. If $|V_S \cap N_F| \leq q + 1$, then $E_S \cap E_2 = \emptyset$, and

$$v(S) = \omega(E_S) = (2q - 1) + (q + 1)|V_S \cap N_F| + (2q + 1)|S \cap N_X|.$$  \hspace{1cm} (3.1)

2. If $|V_S \cap N_F| \geq q + 1$, then $E_S \cap E_1 = \emptyset$, and

$$v(S) = \omega(E_S) = (2q - 1) + (q + 1) + q|V_S \cap N_F| + (2q + 1)|S \cap N_X|.$$  \hspace{1cm} (3.2)

Therefore, for the grand coalition $N$, $v(N) = 6q(q + 1) + q|F|$. Let

$$x(u_i) = q + \frac{3q}{3q + |F|}, \; \forall i = 1, 2, \ldots, |F|;$$

$$x(v_j) = 2q + 1 + \frac{3q}{3q + |F|}, \; \forall j = 1, 2, \ldots, 3q;$$
Membership for Core of LP Games and Other Games

Fig. 2. \( G = (V, E; \omega) \).

\[ N_F = \{ u_i : i = 1, 2, \ldots, |F| \}; \quad N_X = \{ v_j : j = 1, 2, \ldots, 3q \}. \]

(The characters in brackets are the corresponding edge weights)

4 A \( \text{NP}-\text{Hard Total Balancedness Condition} \)

In this section, we prove the problem of testing total balancedness of the Steiner tree game is \( \text{NP}-\text{hard} \). In our proof, we will use 3-PARTITION, which is shown to be \( \text{NP}-\text{complete} \) in strong sense in the book of Garey and Johnson [6].

**Theorem 4.1** Testing total balancedness for Steiner tree games is \( \text{NP}-\text{hard} \).

Proof. Given any instance of 3-PARTITION: A finite set \( \{ a_1, a_2, \ldots, a_{3m} \} \), a bound \( B \in \mathbb{Z}^+ \), and a size: \( s(a) \in \mathbb{Z}^+ \) for each \( a \in A \), such that

\[
\begin{align*}
\frac{B}{3} < s(a) < \frac{B}{2}, & \quad \forall a \in A, \\
\sum_{a \in A} s(a) = mB.
\end{align*}
\]
We let $S = \{S : S \text{ is a 3-set of } A\}$, $p = |S| = C_3^3m$, and $Q = \{Q : Q \text{ is a } (3m-1)\text{-set of } A\}$, $q = |Q| = 3m$.

![Diagram of graph G]

(The quantities in the brackets are the corresponding edge weights.)

**Fig. 3.** $G = (V, E; \omega)$

I. First we construct a weighted graph $G = (V, E; \omega)$.

- The node set $V = \{v_0\} \cup N \cup M$ is given as follows:
  - $\{v_0\}$ is corresponding to the supplier;
  - $N = A$ is corresponding to the set of consumers;
  - $M = S \cup Q \cup \{u_0\}$ is corresponding to the set of switches.

- The edge set $E = E_1 \cup E_2 \cup E_3 \cup E_4 \cup E_5$, and the edge weight function $\omega : E \rightarrow R^+$ are defined as follows:
  - $E_1 = \{(a, S) : a \in S, \text{ for all } a \in A, S \in S\}, \forall e \in E_1, \omega(e) = 6mB$;
  - $E_2 = \{(a, Q) : a \in Q, \text{ for all } a \in A, Q \in Q\}, \forall e \in E_2, \omega(e) = 6mB + \frac{5B}{3}$;
  - $E_3 = \{(S, u_0) : S \in S\}$,
    $$\forall e = (S, u_0) \in E_3, \omega(e) = \begin{cases} B & \text{if } \sum_{a \in S} s(a) \leq B, \\ \frac{3B}{2} & \text{if } \sum_{a \in S} s(a) > B; \end{cases}$$
  - $E_4 = \{(Q, v_0) : Q \in Q\}, \forall e \in E_4, \omega(e) = mB - \frac{B}{3}$;
  - $E_5 = \{(u_0, v_0)\}, \omega(u_0, v_0) = 5mB$.

Denote the corresponding Steiner tree game by $\Gamma(G) \equiv (N, v)$. For convenience of comprehension, we give a sketch of the graph $G = (V, E; \omega)$ in Fig.3. Notice that

1. The weight of each edge $e \in E_1 \cup E_2$ is sufficient large so that in any minimum Steiner tree of $G$ w.r.t. $S \cup \{v_0\}$ ($S \subseteq N$), there are exactly $|S|$ edges in $E_1 \cup E_2$. 
2). The construction of edge \((u_0, v_0)\) is to ensure that for any proper subset of \(A\), the vertices in \(S\) are not needed in the corresponding minimum Steiner tree.

II. We show that the cores of all proper subgames of \(\Gamma(G)\) are non-empty. Let \(S\) be any proper subset of \(N\), \(|S| \leq 3m - 1\). Assume that \(T_S = (V_S, E_S)\) is the minimum Steiner tree of \(G\) w.r.t. \(S \cup \{v_0\}\), then it must be the case \(V_S = \{v_0\} \cup S \cup Q\), where the corresponding set \(Q \in Q\) including \(S\). That is, \(E_S\) consists of one edge in \(E_4\) and \(|S|\) edges in \(E_2\),

\[
v(S) = 6mB|S| + \frac{5}{3}B|S| + mB - \frac{1}{3}B := v_S^*.
\]

Let

\[
x(a) = \frac{v_S^*}{|S|} = 6mB + \frac{5B}{3} + \frac{(3m - 1)B}{3|S|}, \quad \forall a \in S,
\]

easy to verify that it is an element in the core of \(\Gamma(G_S)\).

III. We prove that the game \(\Gamma(G)\) has non-empty core if and only if \(A\) has a 3-partition.

Assume that \(T = (V_N, E_N)\) is the minimum Steiner tree of \(G\) w.r.t. \(N \cup \{v_0\}\), then we have

\[V_N = \{v_0\} \cup N \cup \{S_1, \ldots, S_m\},\]

where \(S_1, \ldots, S_m \in S\) and \(\bigcup_{i=1}^m S_i = A\). That is, \(E_N\) consists of the edge \((u_0, v_0)\), \(m\) edges in \(E_3\) and \(3m\) edges in \(E_1\), and

\[
v(N) \leq 6mB \cdot 3m + m\frac{3B}{2} + 5mB = 18m^2B + \frac{13}{2}mB := v_N^*.
\]

In fact,

a). If \(E_N\) does not contain any edges in \(E_1 \cup E_3 \cup E_5\), then it must consist of \(3m\) edges in \(E_2\) and two edges in \(E_4\), and \(v(N) = 2(mB - \frac{B}{3}) + 3m(6mB + \frac{5B}{3}) > v_N^*\), a contradiction.

b). If \(E_N\) contains \(k\) \((0 < k < 3m)\) edges in \(E_2\), then it must contain one edge in \(E_4\), and \(v(N) \geq (mB - \frac{B}{3}) + k(6mB + \frac{5B}{3}) + 5mB - (3m - k)6mB + \lceil \frac{3m-k}{3} \rceil B > v_N^*\), also a contradiction.

If \(A\) has a 3-partition, then

\[
v(N) = 6mB \cdot 3m + mB + 5mB = 18m^2B + 6mB.
\]

Let \(x(a) = 6mB + 2B, \quad \forall a \in N\). By formula (4.1), easy to verify that it is a core member of \(\Gamma(G)\).

If \(A\) has no 3-partition, suppose the core of \(\Gamma(G)\) is not empty and \(x \in \text{Core}(\Gamma(G))\). Since \(x(N) = v(N) > 18m^2B + 6mB\), there must be a \((3m - 1)\)-element subset \(Q^*\) of \(N\) such that

\[
x(Q^*) > \frac{3m - 1}{3m}(18m^2B + 6mB) = (3m - 1)(6mB + 2B).
\]

By the formula (4.1), we have

\[
v(Q^*) = 6mB(3m - 1) + \frac{5}{3}B(3m - 1) + mB - \frac{B}{3} = (3m - 1)(6mB + 2B) < x(Q^*),
\]
which is contrary to our hypothesis that \( x \in \text{Core}(\Gamma(G)) \).

Therefore, the \( \Gamma(G) \) is totally balanced if and only if \( A \) has an equivalent 3-partition. Also the construction of \( \Gamma(G) \) can be carried out in polynomial time, so testing total balancedness of Steiner tree games is \( NP \)-hard.

**References**

Strong Solutions to the Identification Problem

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Abstract. This work proposes two identification algorithms based on some difficult problems. The first scheme is an extremely simple combination of the idea of one-time passwords and the use of public keys. The second proposal allows identification to be implemented without leaking any secret information during the interaction because it combines the two mentioned concepts with a zero-knowledge proof of the possession of a solution to a difficult graph problem through a challenge-response technique.

1 Introduction

Identification schemes are designed to allow one party (the verifier) to gain assurance that the identity of another (the prover) is as declared. The most common technique is the use of secret time-invariant passwords that are shared between both participants. This scheme corresponds to a maximum-disclosure proof and its major security concern is eavesdropping and subsequent replay of the secret information. There are two well-known solutions to this security problem. The weakest of both methods uses one-time shared passwords, whereas the strongest solution consists in applying the idea of challenge-response protocols that allow the prover A (Alice) to convince the verifier B (the system) that she has some information without revealing it. This idea corresponds to minimum-disclosure proofs and has been implemented through several techniques as for example zero-knowledge proofs. This last concept was formalized in 1986 by Goldreich, Micali and Wigderson, [3], who showed that any problem in NP can be given a zero-knowledge proof. This kind of proofs implies zero disclosure, the use of random numbers as challenges and the application of bit commitment schemes. The recent reference [2] provides an extensive review to the area of probabilistic proofs. There we can find assertions such as: ”The development of techniques for the construction of efficient zero-knowledge proof systems is still of interest”, which refers exactly to the subject of this work.

The work [4] proposes a zero-knowledge proof for the solution to the graph isomorphism problem. Here we use the same problem as basis for the demonstration of the knowledge of a solution to other difficult problems, [1]. Note also that in our proposals there is no need to know how to solve any difficult problem because the user builds an adequate graph for each chosen solution.
This paper is organized as follows: In the following section we comment the main ingredients of our protocols. Sections 3 and 4 present respectively a simple protocol based on one-time passwords and a scheme based on the challenge-response technique. Finally, the work ends with some conclusions and open problems.

2 Ingredients

The proposed schemes combine challenge-response techniques with one-time passwords. Now we comment some of the main aspects of these ingredients.

2.1 Challenge-Response

According to the challenge-response technique, the steps followed by the participants A and B can be generally described in the following way:

1. A chooses a random element (witness) of a predefined set, which defines a set of questions (challenges) whose answers (responses) only she knows.
2. B chooses one of the challenges.
3. A provides the corresponding response.
4. B verifies the received response and witness.

If the number of challenges is \( c \), then the probability of A’s fraud after \( m \) iterations is \( c^{-m} \). Thus, to decrease A’s fraud probability, the number \( m \) of iterations should be increased. Also, to protect the secret information, A should answer, at the most, to one question for each witness, and never reuse any previous witness. On the other hand, formal zero-knowledge proofs should verify the properties of completeness, soundness and zero-knowledge, \([2]\).

2.2 Graphs

Another important ingredient for these protocols is the difficult graph problem used as basis to guarantee their security. In order to make more efficient the verification phase, this problem should be ”one-way”, that is to say, it should be such that it is difficult to find a solution but however it is very simple to verify a solution. In these proposals the user has to follow a building process for the construction of an adequate graph from a randomly chosen solution. This process has a different computational complexity depending on the specific graph problem, so this aspect has to be also considered when selecting the graph problem. So, it seems that the most suitable difficult graph problems are those where the solution is a vertex subset, such as the vertex cover, the independent set or the clique problems. Regarding the building process, in both protocols, once chosen a problem \( P \), a number of vertices \( n \) and a solution \( S \), the user A has to build a graph \( G \) where \( S \) is a solution to the problem \( P \). In order to do it, A can apply one of the two following methods:

- **CM)** Constructive method: By adding edges to the null graph \( N_n \)
- **DM)** Destructive method: By deleting edges from the complete graph \( K_n \)
2.3 Representations

Throughout this work, every secret identification password corresponds to a solution of a difficult graph problem. So, regarding memory requirements, special attention should be paid to the solutions, graphs, functions and problems representations. Every solution is here symbolized by the decimal value of the binary representation where the unitary bits indicate the presence of the corresponding elements of the solution, so in the first protocol, since this is the only object that has to be sent from A to B, the communication complexity is not a great problem because it only implies to send a binary vector whose length depends on the kind of the used problem. On the other hand, the problems have to be published in a certified data base. So, a code dictionary for the list of possible problems can be published in the same directory allowing the use of a short code corresponding to each problem.

3 Non-probabilistic Protocol

Some of the main ideas of the scheme proposed in the following section are introduced now with the definition of the first extremely simple proposal based on one-time passwords. The description of this protocol is as follows:

Precalculación
1. A chooses a graph problem P and an adequate solution, and states the integer S corresponding to this solution
2. A builds a corresponding graph G for P and S
3. A states as her public key the pair (P,G)

Interaction
One-time secret password: A positive integer S
One-time public information: A graph G and a problem P
1. A sends to B the solution S
2. B checks that S is a valid solution according to A’s public key (P,G)

This scheme is non probabilistic, so after applying it, there should be no doubt about the identity of the user A thanks to the difficulty of the problem and the well choice of the parameters. Another advantage of this proposal is the no need to share any secret information between both parties, in contrast with other well-known one-time password schemes. Note that with this protocol, once used a solution S, it loses its secret condition, so it is necessary to update it after each identification process. So, during the precalculation phase, A has to carry a building step that is completely independent from previous performances of the identification scheme because the solution changes in every running. In order to make more efficient this building step, in the following section we will propose a probabilistic proof where the user A can take approach of previous performances to build new graphs.
4 Probabilistic Protocol

In this section we present a general interactive challenge-response scheme based on difficult graph problems. Now an informal description is given prior to the formal definition. The protocol can be divided into two parts. The first part is an off-line precalculation of the prover where she chooses her fixed secret identification number $S$, and her one-time public information: a problem $P$ and a graph $G$. The second part corresponds to the zero-knowledge procedure composed of $m$ iterations where $A$ commits to certain information, $B$ chooses a challenge, $A$ responses and $B$ verifies the answer and the commitment, so that finally it accepts that $A$ possesses a valid secret identification number. So, the formal description of the algorithm is as follows:

Precalculation

1. $A$ chooses a graph problem $P$ and a suitable solution, and states the integer $S$ corresponding to this solution
2. $A$ builds an adequate graph $G$ where $S$ represents a solution to $P$
3. $A$ updates the pair $(P,G)$ in the public data base

Interaction

Fixed secret information: A positive integer $S$

One-time public information: A graph $G$ and a problem $P$

The following four steps are executed $m$ times:

1. $A$ generates at random an isomorphic copy $H$ of $G$, and sends it to $B$ jointly with the cipher of both, the isomorphism $f$ and the solution $T=f(S)$ to the problem $P$ in $H$
2. $B$ obtains from the public data base the information $(P,G)$ associated to the user $A$ and selects randomly a bit $b$. Depending on its value, it requests her to decipher:
   (a) a characteristic question of the solution $T$, if $b=0$
   (b) the isomorphism $f$, if $b=1$
3. $A$ responses to the challenge
4. Depending on the selected challenge, $B$ verifies the corresponding cipher and checks that:
   (a) $T$ fulfills the properties of a valid solution to the problem $P$ in the graph $H$, if $b=0$
   (b) $f$ transforms the graph $G$ in its isomorphic copy $H$, if $b=1$

In this scheme there exists certain probability that a fraudulent prover passes the verification. On the other hand, the proof of its completeness, soundness and computational zero-knowledge properties is easily affordable. Note also that with this scheme the solution $S$ does not loose its secret condition, so it is not necessary to change it after each running and there is no need to share any secret information. Furthermore, this probabilistic proof has the advantage of requiring less off-line operations than the previous proposal, provided that the building process can be implemented in a way that for one fixed solution and

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1 The step 2(a) can be replaced by the total disclosure of the solution $T$ when the knowledge of its cardinality does not help to deduce $S$. 

for each different problem, the generation of every graph can take approach of
previous generations, or even though it can be easily carried out in parallel.
Finally, an important aspect of this protocol is that it requires the assumption
of the following hypothesis:
"The knowledge of different pairs of difficult problems and corresponding
graphs for which there exists a common solution does not help to solve any of
the graph problems".

5 Conclusions and Open Problems

The primary aim of this work is the design of cryptographic protocols in order
to solve the identification problem. The schemes here proposed constitute two
elegant models that allow the application of many different problems from the
Graph Theory. They are the result of the combination of the idea of one-time
passwords with public keys in the first case, and commitment schemes with
challenge-response in the second. Also the NP-completeness is suggested to be
used here with a positive use as a basis for both schemes.

Some questions about the use of the protocols to be recalled are the following:

a) the usefulness of these schemes relies on the great variety of graph problems
that can be used as basis for the application of both protocols.

b) the basis problem should be "one-way", that is to say, such that it is
difficult to find a solution but however it is extremely simple to verify a solution.

A forthcoming version of this work will include the comparison among the
computational efficiency of the algorithms, the search of the most suitable in-
stance graphs and problems, the analysis of minimal assumptions for the imple-
mentations and the study of the possible parallelization of the second proposal.

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Area Efficient Exponentiation
Using Modular Multiplier/Squarer in GF($2^m$)

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Abstract. This paper presents a new exponentiation architecture and multiplier/squarer for GF($2^m$), which uses a standard basis representation. The proposed multiplier/squarer is used as kernel architecture of exponentiation. Although the proposed multiplier/squarer computes the multiplication and squaring operations at the same time in GF($2^m$), the common parts existing in both operations are only executed once, thereby reducing the required hardware compared to related systolic circuits. The proposed multiplier/squarer can be easily applied to exponentiation architecture. It is also well suited to VLSI implementation because of its regularity, modularity, and unidirectional data flow.

1 Introduction

Finite fields GF($2^m$) have found useful in error-correcting codes and cryptography [1][2], however, these applications usually require the computation of multiplications, squares, and exponentiations in GF($2^m$). Modular exponentiation is the core computation of numerous public key cryptosystems including RSA. As such, to satisfy real-time requirements, efficient hardware structures for such operations are desirable.

The conventional method for computing squares in GF($2^m$) over a standard basis involves a multiplier. Recently, a low latency squarer for GF($2^m$) with bi-directional data flow over a standard basis was proposed which involves less chip area and a smaller latency than using a multiplier in GF($2^m$). For systolic array implementations of exponentiation architecture for GF($2^m$), the architectures in [3] and [8] can be used. The circuit in [3] consists of the $2(m-1)$ multipliers in [5]; the circuit in [4] requires the $m-1$ power-sum circuits. The architecture in [3] involves bi-directional data flow, while those in [6] and [7] involve unidirectional data flow.

This paper proposes a new systolic multiplier/squarer and new systolic exponentiation architecture for GF($2^m$) with unidirectional data flow over a standard basis. Both circuits involve a smaller latency and less hardware complexity than

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existing related systems, plus they are well suited to VLSI implementation with fault-tolerant design.

## 2 Exponentiation Architecture for GF($2^m$)

Unlike addition and multiplication, exponentiation in GF($2^m$) is very similar to an operation for ordinary integers. Let $A(x)$ and $B(x)$ be elements in GF($2^m$), and $P(x)$ be the primitive irreducible polynomial of degree $m$ [5]. The three polynomials $A(x)$, $B(x)$ and $P(x)$ are written as

$$A(x) = \sum_{k=1}^{m-1} a_k x^k, \quad B(x) = \sum_{k=1}^{m-1} b_k x^k, \quad P(x) = \sum_{k=1}^{m} p_k x^k, \quad p_m = 1$$

If $e = [e_{m-1}, e_{m-2}, \ldots, e_0]$ is the exponent, a right-to-left algorithm can be performed to compute $A(x)^e \mod P(x)$ over GF($2^m$). Ordinary right-to-left algorithm in [8] can be slightly changed to Alg. 1, which uses a new multiplication/squaring algorithm, M/S algorithm, as a sub-function for exponentiation algorithm. M/S algorithm will be described in section 3.

[Algorithm 1] Exponentiation Algorithm using M/S Alg.

**Input** $A(x)$, $e$, $P(x)$

**Output** $E(x) = A(x) \mod P(x)$

**Step 1** $E(x) = 1$, $T(x) = A(x)$

**Step 2** for $i = 0$ to $m-1$ do

**Step 3** if $e_i = 1$ ($E(i-1)(x), T(i-1)(x)$) = M/S ($E(i-1)(x), T(i-1)(x)$)

else ($E(i)(x) = E(i-1)(x), T(i)(x)$) = M/S ($E(i-1)(x), T(i-1)(x)$)

In step 3, M/S algorithm is called with two parameters $E^{(i-1)}(x)$ and $T^{(i-1)}(x)$, and it returns two computation results of multiplication, $E^{(i-1)}(x)*T^{(i-1)}(x)$, and squaring,

![Fig. 1. Proposed exponentiation architecture in GF($2^m$)](image_url)
\( T^{(i-1)}(x) \ast T^{(i-1)}(x) \). Square result is stored to \( T^{(i)}(x) \). But, the variable \( E^{(i)}(x) \) stores a value depends on the value of \( e_i \). Returned multiplication result is passed to \( E^{(i)}(x) \) \( e_i \) has value 1, but previous result, \( E^{(i-1)}(x) \), is stored to \( E^{(i)}(x) \) in the other case.

Fig. 1 shows a new sequential exponentiation architecture using a new systolic multiplier/squarer, which operates Alg. 1. This architecture includes a multiplier/squarer as its kernel architecture. The multiplier/squarer consists of main two parts, the MUL and the SQR. The MUL operates a multiplication operation whereas the SQR operates a square operation. IPL register keeps an irreducible polynomial. While the multiplier/squarer computes the multiplication operation, it also processes the squaring operation concurrently forming the next square term. The switch depicted in Fig. 1 decides the multiplication result, which chooses the previous result or the current result depends on the exponent.

3 Systolic Multiplier/Squarer for GF\((2^m)\)

The computation steps for the modular multiplication \( M(x) = A(x)B(x) \mod P(x) \) and square \( S(x) = A(x)A(x) \mod P(x) \) are follows as:

\[
\begin{align*}
M(x) &= A(x)B(x) \mod P(x) \\
&= b_0A(x) + b_1[A(x)x \mod P(x)] + \ldots + b_{m-1}[A(x)x^{m-1} \mod P(x)] \\
S(x) &= A(x)A(x) \mod P(x) \\
&= a_0A(x) + a_1[A(x)x \mod P(x)] + \ldots + a_{m-1}[A(x)x^{m-1} \mod P(x)]
\end{align*}
\]

In the above two equations (1) and (2), \([ ]\) is common to both. Thus, it is much more efficient to compute the common parts only once when multiplication and squaring operations are carried out simultaneously. As such, equation (1) can be reformulated into the recurrence form, called LSB-first form as follows: for \( 1 \leq i \leq m \)

\[
M^{(i)}(x) = M^{(i-1)}(x) + b_{i-1}A^{(i-1)}(x), A^{(0)}(x) = A^{(i-1)}(x) \mod P(x)
\]

where \( A^{(0)}(x) = A(x), M^{(0)}(x) = 0, \) and \( M^{(i)}(x) = b_0A(x) + b_1[A(x)x \mod P(x)] + b_2[A(x)x^2 \mod P(x)] + \ldots + b_{i-1}[A(x)x^{i-1} \mod P(x)] \). For \( i=m, M^{(m)}(x) = M(x) = A(x)B(x) \mod P(x) \). Note that the two equations in (3) can be performed in parallel.

The above squaring operation (2) can be reformulated to the following LSB-first recurrence form just like equation (3) derived for multiplication: \( 1 \leq i \leq m \)

\[
S^{(i)}(x) = S^{(i-1)}(x) + a_{i-1}A^{(i-1)}(x), A^{(0)}(x) = A^{(i-1)}(x) \mod P(x)
\]

where \( A^{(0)}(x) = A(x), S^{(0)}(x) = 0, \) and \( S^{(i)}(x) = a_0A(x) + a_1[A(x)x \mod P(x)] + a_2[A(x)x^2 \mod P(x)] + \ldots + a_{i-1}[A(x)x^{i-1} \mod P(x)] \). For \( i=m, S^{(m)}(x) = S(x) = A(x)A(x) \mod P(x) \). Note that the two equations in (4) can also be performed in parallel.

The following bit-wise LSB-first algorithm computing both modular multiplication and squaring simultaneously can be derived from the above equations (3) and (4):

[Algorithm 2] Multiplication and Squaring Algorithm:

\[
M/S(E(x), T(x))
\]

Input \( A(x)=T(x)=(a_{n-1}, a_{n-2}, \ldots, a_0) \)

\( B(x)=E(x)=(b_{n-1}, b_{n-2}, \ldots, b_0) \)
$P(x) = (1, P_{m-1}, P_{m-2}, \ldots, P_0)$

Output
$M(x) = A(x) B(x) \mod P(x)$, $S(x) = A(x) A(x) \mod P(x)$

Step 1
$M^{(0)} = (M^{(0)}_{m-1}, M^{(0)}_{m-2}, \ldots, M^{(0)}_0) = (0, 0, \ldots, 0)$
$S^{(0)} = (S^{(0)}_{m-1}, S^{(0)}_{m-2}, \ldots, S^{(0)}_0) = (0, 0, \ldots, 0)$
$A^{(0)} = (A^{(0)}_{m-1}, A^{(0)}_{m-2}, \ldots, A^{(0)}_0) = (a_{m-1}, a_{m-2}, \ldots, a_0)$
$A^{(i-1)} = 0, 0 \leq i \leq m$

Step 2 for $i = 1$ to $m$
Step 3 for $j = 1$ to $m$
Step 4
$M^{(i)}_{m-j} = M^{(i-1)}_{m-j} + B_{m-j} A^{(i-1)}_{(m-j)}$

Step 5
$S^{(i)}_{m-j} = S^{(i-1)}_{m-j} + A^{(i-1)}_{m-j} + p_{m-j} A^{(i-1)}_{m-j}$

Step 6
$A^{(i)}_{m-j} = A^{(i-1)}_{m-j-1} + p_{m-j} A^{(i-1)}_{m-j}$

Fig. 2 shows a DG of Alg. 2 over GF($2^4$). The node at position $(i,j)$, $(1 \leq i, j \leq m, m = 4)$ performs three recurrence equations of Alg. 2 in parallel. DG with regular data flows can be easily transformed to a systolic array via a space-time transformation [4].

The time transformation vector $t$ is equal to $[2, 1]$ because the vector of the equitemporal hyperplane is $[1, -2]$. The minimum number of PEs to be mapped is when the $j$ column, $[0, 1]$ can be fixed to the projection vector. The space transformation vector $s$, which is vertical to the projection vector, is $[1, 0]$. By using the vector $s$, the nodes on the same row in Fig. 2 can be mapped to the same PE. The proposed systolic multiplier/squarer is shown in Fig. 3.

### 4 Conclusion

This paper have presented a new systolic multiplier/squarer computing multiplication and squaring operations at the same time, and a new serial-in-serial-out exponentiation architecture for GF($2^m$) using a standard basis representation. Although computing multiplication and squaring operations at the same time in GF($2^m$), the proposed systolic squarer/multiplier requires less hardware than a dedicated systolic multiplier.
Table 1. Comparison of architectures on GF($2^m$)

<table>
<thead>
<tr>
<th></th>
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<tbody>
<tr>
<td>No. of PEs</td>
<td>$m$</td>
<td>$m$</td>
<td>$m$</td>
</tr>
<tr>
<td>Latency</td>
<td>$3m$</td>
<td>$3m$</td>
<td>$3m$</td>
</tr>
<tr>
<td>PE complexity</td>
<td>3 AND gates</td>
<td>2 AND gates</td>
<td>3 AND gates</td>
</tr>
<tr>
<td></td>
<td>2 XOR gates</td>
<td>2 XOR gates</td>
<td>3 XOR gates</td>
</tr>
<tr>
<td></td>
<td>2 MUX(1X2)</td>
<td>2 MUX(1X2)</td>
<td>3 MUX(1X2)</td>
</tr>
<tr>
<td></td>
<td>10 Latches</td>
<td>11 Latches</td>
<td>14 Latches</td>
</tr>
<tr>
<td>Critical path</td>
<td>1AND+2XOR+</td>
<td>1AND+1XOR+</td>
<td>1AND+1XOR+</td>
</tr>
<tr>
<td></td>
<td>1MUX+1Latch</td>
<td>1MUX+1Latch</td>
<td>1MUX+1Latch</td>
</tr>
</tbody>
</table>

As shown in Table 1, the PE complexity of the proposed multiplier/squarer was higher than that of existing multipliers. This fact is undeniable because the proposed multiplier/squarer computes modular multiplication and squaring simultaneously, whereas the other multipliers compute only modular multiplication. The proposed hardware is well suited to VLSI implementation because of its regularity, modularity, and unidirectional data flow.

References

A Space Saving Trick for Directed Dynamic Transitive Closure and Shortest Path Algorithms

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Abstract. We present a simple space saving trick that applies to many previous algorithms for transitive closure and shortest paths in dynamic directed graphs. In these problems, an update can change all edges incident to a node. The basic queries on reachability and distances should be answered in constant time, but also paths should be produced in time proportional to their length. For:

Transitive closure of Demetrescu and Italiano (FOCS 2000)
Space reduction from $O(n^3)$ to $O(n^2)$, preserving an amortized update time of $O(n^2)$.

Exact all-pairs shortest dipaths of King (FOCS 1999)
Space reduction from $\tilde{O}(n^3)$ to $\tilde{O}(n^2\sqrt{n}\delta)$, preserving an amortized update time of $\tilde{O}(n^2\sqrt{n}\delta)$, where $\delta$ is the maximal edge weight.

Approximate all-pairs shortest dipaths of King (FOCS 1999)
Space reduction from $\tilde{O}(n^3)$ to $\tilde{O}(n^2)$, preserving an amortized update time of $\tilde{O}(n^2)$.

Several authors (Demetrescu and Italiano, FOCS 2000, and Brown and King, Oberwolfach 2000) had discovered techniques to give a corresponding space reduction, but these techniques could be used to show only the existence of a desired dipath, and could not be used to produce the actual path.

1 Introduction

This paper is on saving space for dynamic transitive closure and shortest paths in dynamic directed graphs. For example, the space saving for transitive closure is from $O(n^3)$ to $O(n^2)$. Papers on algorithm are often more concerned with time than space, but a bad space complexity hits like a wall, whereas time only kills slowly. Ignoring for a moment the $O$-notation, if a computer has gigabyte of memory, with $n^3$ space, you can only run on graphs with up to 1,000 nodes, whereas with $n^2$ space, you can deal with more than 30,000 nodes. Our space savings techniques are simplifying and tend to decrease rather than increase the constants hidden in the $O$-notation, both in time and in space.
A fully dynamic graph algorithm is a data structure for a graph which implements an on-line sequence of update operations that insert and delete edges in the graph and answers queries about a given property of the graph. A dynamic graph algorithm should process queries quickly and must perform update operations faster than computing from scratch (as performed by the fastest “static” algorithm).

In this paper, we give dynamic algorithms for weighted directed graphs. The problems considered are as follows, where the queries range over all pairs of vertices $u$ and $w$:

**Transitive closure:**
Reachability query: Is there a path from $u$ to $w$ in the current graph?
Path query: Return a path from $u$ to $w$.

**Exact all-pairs shortest paths:**
Distance query: What is the distance from $u$ to $w$?
Path query: Return a path from $u$ to $w$.

**Approximate all-pairs shortest paths:**
Distance query: What is an upper bound on the distance from $u$ to $w$ within a factor of $(1 + \epsilon)$?
Path query: Return a path from $u$ to $w$ of length within a factor $(1 + \epsilon)$ of the distance.

The following update operations are allowed:

- multi-insert($E_v$): inserts a set of edges incident to the same vertex $v$.
- delete($e$): deletes any edge currently in the graph.

In this paper, we are only considering amortized time bounds. Deleting an edge is free in that it is paid for by the preceding insertion/initialization of the edge.

We present a simple space saving trick that for all the above problems reduce the space substantially in the current fastest algorithm. More precisely, assuming that the graphs start with an empty edge set, we achieve:

**Transitive closure of Demetrescu and Italiano** [3,4] Space reduction from $O(n^3)$ to $O(n^2)$, preserving an amortized multi-insert time of $O(n^2)$.

**Exact all-pairs shortest dipath of King** [9] Space reduction from $\tilde{O}(n^3)$ to $\tilde{O}(n^2 \sqrt{n\hat{b}})$, preserving an amortized multi-insert time of $\tilde{O}(n^2 \sqrt{n\hat{b}})$, where $\hat{b}$ the maximal edge weight.

**Approximate all-pairs shortest dipaths of King** [9] Space reduction from $\tilde{O}(n^3)$ to $\tilde{O}(n^2)$, preserving an amortized multi-insert time of $\tilde{O}(n^2)$.

The above update times should be compared with the best static algorithms. Transitive closure and approximate shortest paths can be computed essentially as fast matrix multiplication [8,10,13], which takes $\tilde{O}(n^{2.376})$ time, and for exact all pairs shortest paths, can be computed in $\tilde{O}(n^{2.575})$ time, again using fast
matrix multiplication. Thus, the amortized update times may not seem that impressive. However, fast matrix multiplication is considered rather slow, and it is an open problem to construct a combinatorial algorithm just for transitive closure with sub-cubic running time. Our algorithms are purely combinatorial, and in this view, the dynamic speed-up is optimal in that we can construct a graph from scratch with a multi-insert from each vertex, paying a total cost of $n \times O(n^2) = O(n^3)$ to construct the transitive closure, thus matching the best static combinatorial time bound. Our reduction of the space from $O(n^3)$ to $O(n^2)$ adds optimality with respect to space.

Our ideas also apply to the older single source algorithms of Even and Shiloach [5] and of Ramalingam and Reps [11], improving their internal space from $O(m)$ to $O(n)$. Hence, if used for all pairs shortest paths, the improvement is from $O(mn)$ to $O(n^2)$. We are here particularly interested in Ramalingam and Reps’ algorithms because it has proved efficient in practice [7,6]. Our idea also simplifies the implementation which typically reduces the worked performed.

A common theme of all the above algorithms is that they keep tables of distance information along with witnesses. A table entry can only increase if all its witnesses are lost. This theme goes back to Even and Shiloach’s decremental single shortest path algorithm [5], and is also used in the fully-dynamic single source shortest path algorithm of Ramalingam and Reps [11]. Unfortunately, there may be many witnesses for each entry, and it is the storage of these witnesses that cause a prohibitive space-overhead.

Recently space improvement similar to ours were reported by Brown and King [2] and of Demetrescu and Italiano [4]. Their idea was to maintain a counter of the number of witnesses for each entry. Again, the entry can only increase if its counter goes to 0. The problem in their solution is that they loose track of witnesses, and then they cannot answer path queries.

Our challenge of identifying witnesses rather than just knowing their existence is similar in spirit and motivation to that of finding witnesses for Boolean matrix multiplication as done by Alon et al. [1], though the techniques we apply here are combinatorial and completely different.

Our contribution is to observe that if we are careful about the order in which we scan for witnesses, then we only need to store the first witness that we meet. When a witness is lost, we just continue the scan for the next witness. Our problem is then to identify scanning orders and properly define witnesses so that the same potential witnesses are not scanned repeatedly.

Contents First, in §2 we present the simplest version of the trick, applying it to Even and Shiloach’s classic algorithm [5]. In §2 we also point out the modification needed for Ramalingam and Reps’ algorithm [11]. Next, in §3 we show how to tailor the trick to King’s exact shortest paths and transitive closure algorithms [9]. For space reasons we defer the treatment of King’s exact and approximate shortest paths and Ramalingam and Reps shortest paths algorithm to the journal version of this paper.
Notation We work on a dynamic directed weighted graph \( G = (E, V) \) with edge weight function \( \ell : E \to \mathbb{N} \). For any subgraph \( H \) of \( G \), \( V(H) \) is its vertex set and \( E(H) \) its edge set. For any pair of vertices \( (v, w) \), \( \text{dist}(v, w) \) denotes their distance in \( G \). For each vertex \( v \), \( \text{in}(v) \) and \( \text{out}(v) \) denotes its incoming and outgoing edges in \( G \).

2 Even and Shiloach’s Algorithm

Our basic idea is most easily presented in terms of Even and Shiloach’s decremental single shortest path algorithm [5]. Their result is that we can maintain distances up to some threshold \( \Delta \) in \( O(m\Delta) \) total time. In [5] they just consider unweighted graphs and threshold distance \( \Delta = n \), giving them an \( O(mn) \) bound for maintaining all distances from the source. We generalize this to weighted graphs.

For a given source node \( s \), the Even-Shiloach algorithm maintains a shortest path graph \( H_s \) which is the union of all shortest paths from \( s \) of length at most \( \Delta \). For each vertex \( v \), we maintain its distance \( d_s(v) \) from \( s \), and, for each distance \( i = 0, ..., \Delta \), we maintain the set of vertices \( v \) with \( d_s(v) = i \). During our response to an edge deletion, \( d_s(v) \) will be incremented by one at a time until it reaches the new distance from \( s \) to \( v \).

In the shortest path graph \( H_s \), each edge \((u, v)\) is a witness of the current distance to \( v \), in the sense that \( \text{dist}(s, v) = d_s(v) = d_s(u) + \ell(u, v) = \text{dist}(s, u) + \ell(u, v) \). Hence, when an edge \((u, v)\) is deleted, it affects distances only if \((u, v)\) is in \( H_s \), and there is no other edge \((x, v)\) to \( v \) in \( H_s \).

Suppose the deleted edge \((u, v)\) is a last witness to \( v \) in \( H_s \). This implies that \( d_s(v) \) has increased, and the outgoing edges \((v, w)\) are no longer valid witnesses of \( d_s(w) \). A cascading effect is started: when a node \( w \) loses all its incoming edges in \( H_s \), we must delete all its outgoing edges in \( H_s \).

To update \( H_s \) efficiently, the cascade is performed one distance \( i \) at the time. Let \((u, v)\) be the original edge deleted. Starting with \( i \leftarrow d_s(v) \), we do as follows. For each vertex \( w \) with \( d_s(w) = i \) and no incoming edges in \( H_s \), we remove all its outgoing edges in \( H_s \), and increment \( d_s(w) \) by one. This preserves \( d_s(w) \leq \text{dist}(s, w) \), and now, for all \( u \) with \( d_s(u) \leq i \), we know that \( d_s(u) = \text{dist}(s, u) \). Next, for each \( w \) with \( d_s(w) \) increased, we consider all its incoming edges \((x, w)\) in \( G \). The edge \((x, w)\) witnesses \( d_s(w) = \text{dist}(s, w) \) if \( d_s(x) \leq i \) and \( d_s(x) + \ell(x, v) = d_s(w) \), and if so, \((x, w)\) is added to \( H_s \). If no such witness is found, we still have \( d_s(w) < \text{dist}(s, v) \). We now set \( i \leftarrow i + 1 \) and repeat. We terminate when \( i = \Delta + 1 \) or when we get to a distance \( i \) where all nodes have incoming edges in \( H_s \).

For a detailed description and analysis of the above algorithm, the reader is referred to [5]. The running time is dominated by the scanning for witnesses; each time we increment \( d_s(w) \), we scan all its incoming edge to see if they witness the new distance. Since this can happen at most \( \Delta \) times for each \( w \), the total cost is \( O(\Delta \cdot \sum_{w \in V(G)} |\{(x, w) \in E(G)\}|) = O(\Delta m) \).
2.1 Our Space Saving Trick

The internal space of the Even-Shiloach algorithm, i.e. the space beyond the input graph $G$, is $O(m)$ for storing the shortest path graph $H_s$. We will reduce this to $O(n)$. If we want a decremental data structure for each node, then the space is reduced from $O(mn)$ to $O(n^2)$.

For each vertex $v$, its list $\text{in}(v)$ of incoming edges in $G$ stores the incoming edges in some order. Instead of storing all of $H_s$, for each $v$, we store only the edge $(x,v)$ from $H_s$ which is first in $\text{in}(v)$. We then have the unique shortest path tree $T_s$ from $s$, spanning all nodes within distance $\Delta$ and such that the parent pointer of each node $v$ is the first edge in $\text{in}(v)$ which is on a shortest path from the source to $v$.

The deletion of an edge $(u,v)$ only has a consequence if $(u,v)$ is in $T_s$. In that case, we search for a replacement witness. Starting from the successor of $(u,v)$, we scan the edges in $\text{in}(v)$, stopping if we reach a new witness $(x,v)$ with $d_s(x) + \ell(x,v) = d_s(v)$. If that happens, we replace $(u,v)$ with $(x,v)$ in $T_s$, and stop. Else we conclude that $v$ has also lost its last incoming edge in $H_s$. As in the Even-Shiloach algorithm, each incoming edge to each vertex $v$ is considered at most once for each value of $d_s(v)$. Hence, we preserve the running time of $O(\Delta m)$. However, since $T_s$ is a tree, we have reduced the internal space from $O(m)$ to $O(n)$. Also we save work when edges from $H_s \setminus T_s$ are deleted.

3 King’s Algorithms

As the basis for all the fully-dynamic algorithms of King in [9], we want to maintain all distances up to some threshold $\Delta$. Besides edge deletions, we have multi-inserts that for each vertex $v$ can add any set of edges incident to $v$. The amortized cost per multi-insert is $O(\Delta n^2)$. This also pays for all deletions. The space is $O(n^3)$, but we reduce it to $O(\Delta n^2)$.

The algorithm uses the Even-Shiloach algorithm to maintains $2n$ deletions-only data structures for distances up to $\Delta$. For each vertex $v$, we have one $\text{Out}(v)$ for distances from $v$, i.e. with $v$ as source, and one $\text{In}(v)$ with distances to $v$, i.e. with $v$ as a sink. For each pair $(u,w)$ of vertices, $\text{dist}_v(u,w)$ is the distance from $u$ to $v$ in $\text{In}(v)$ plus the distance from $v$ to $w$ in $\text{Out}(v)$.

We start with a graph with no edges. Whenever we make a multi-insert of edges incident to $v$, we reinitialize the pair $(\text{In}(v), \text{Out}(v))$ at amortized cost $O(\Delta m)$. When an edge is deleted, it is deleted from all pairs $(\text{In}(v), \text{Out}(v))$ created after it was inserted.

**Lemma 1.** If the distance from $u$ to $w$ is at most $\Delta$, $\text{dist}(u,w) = \min_v \text{dist}_v(u,w)$.

**Proof.** Obviously, $\text{dist}(u,w) \leq \text{dist}_v(u,w)$ for all $v$. Consider an arbitrary shortest path from $u$ to $w$ and let $v$ be the last vertex updated on this path. Then all the edges of this path are contained in $(\text{In}(v), \text{Out}(v))$ so $\text{dist}_v(u,w) = \text{dist}(u,w)$. 

Thus, our problem is to keep track of the vertex \( v \) minimizing \( \text{dist}_v(u, w) \) for each pair \((u, w)\) of vertices. King’s idea is, for each possible distance \( d \in \{1, \ldots, \Delta\} \), to maintain the list \( L(u, w, d) \) of vertices \( v \) with \( \text{dist}_v(u, w) = d \).

The King algorithm: Whenever a vertex \( u \) increases its distance to \( v \) in \( \text{In}(v) \), this increases all distances \( \text{dist}_v(u, w) \leq \Delta \). If \( d \) and \( d' \) are the old and the new distance of \( v \), \( v \) has to be moved from \( L(u, w, d) \) to \( L(u, w, d') \). This move takes constant time. Each vertex \( u \) increases its distance at most \( \Delta \) times in \( \text{In}(v) \), and each time, we have to perform a move between lists for each vertex \( w \) with \( \text{dist}_v(u, w) \leq \Delta \). Thus, \( u \) pays a total cost of \( O(\Delta n) \) for increases to its distance in \( \text{In}(v) \). The cost for increases to its distance in \( \text{Out}(v) \) is the same. Thus, for each pair \((\text{In}(v), \text{Out}(v))\), the total cost of distance increases is \( O(\Delta n^2) \), which is attributed to the multi-insert creating \((\text{In}(v),\text{Out}(v))\).

To get the distance matrix, for each \((u, v, w)\), we wish to reduce the space \( O(n) \) using the trick from [2,1]. However, that trick assumed for each vertex \( w \), that we could traverse \( \text{in}(w) \) in some fixed order. That is, if \((x, w)\) was the witness of \( w \) in \( T_v \), and if \((x, w)\) was deleted, we could continue the traversal for witnesses from the successor of \((x, w)\) in \( \text{in}(w) \) knowing that the same edge would never be considered twice for the same distance \( d_v(w) \).

To deal with the differing incidence lists, we spend \( O(\Delta n^2) \), rather than \( O(\Delta m) \), to maintain each pair \((\text{In}(v), \text{Out}(v))\). This does not affect our asymptotic cost, since we have already incurred this cost for each multi-insert. Hence we can spend \( O(n) \) time to traverse \( \text{in}(w) \) in \( \text{Out}(v) \).

We fix one arbitrary global ordering \( \prec \) of all the vertices. For each vertex \( w \) its list \( \text{in}(w) \) in \( \text{Out}(v) \) will be ordered as \( \prec \) orders its predecessors. However, \( \text{in}(w) \) will be stored only implicitly. We keep a global \( n \times n \) incidence matrix which for each \((x, y)\) \( \in V^2 \) which stores its length \( \ell(x, y) \) or \( \infty \) if \((x, y)\) is not an edge.

3.1 Improving the Space

To reduce the space of the above data structure to \( O(\Delta n^2) \), we need to address both the pairs \((\text{In}(v), \text{Out}(v))\) and the witness lists \( L(u, w, d) \).

\((\text{In}(v), \text{Out}(v))\): For each pair \((\text{In}(v), \text{Out}(v))\), we wish to reduce the space to \( O(n) \) using the trick from [2,1]. However, that trick assumed for each vertex \( w \), that we could traverse \( \text{in}(w) \) in some fixed order. That is, if \((x, w)\) was the witness of \( w \) in \( T_v \), and if \((x, w)\) was deleted, we could continue the traversal for witnesses from the successor of \((x, w)\) in \( \text{in}(w) \) knowing that the same edge would never be considered twice for the same distance \( d_v(w) \).

To deal with the differing incidence lists, we spend \( O(\Delta n^2) \), rather than \( O(\Delta n^2) \), to maintain each pair \((\text{In}(v), \text{Out}(v))\). This does not affect our asymptotic cost, since we have already incurred this cost for each multi-insert. Hence we can spend \( O(n) \) time to traverse \( \text{in}(w) \) in \( \text{Out}(v) \).
We traverse \( in(w) \) in \( Out(v) \) by going through all the vertices \( x \in V \) in the order \( \prec \) and keeping a pointer to our current location in the order. For each \( x \), we check if \( (x, w) \) is an appropriate witness, i.e., if \( dist_v(v, x) + \ell(x, w) = dist_v(v, w) \). For each edge \( (x, w) \) which serves as a witness, we keep a pointer back to \( v \), so that when \( (x, w) \) or \( dist_v(v, x) \) is changed, we know to look for another witness.

We only increment \( dist_v(v, w) \) after the order is traversed. We only decrement \( dist_v(v, w) \) when \( (In(v), Out(v)) \) is reinitialized. At both times, we reset the pointer to the beginning of the order. Note that in this space-saving version, we may inadvertently add a witness which was created after the last time \( Out(v) \) was initialized.

We maintain a shortest path tree of edge witnesses as described in Section 2.1, so that nodes affected by a distance increase can be easily determined.

As claimed, the total time to traverse the order is \( O(n) \), giving the amortized cost of \( O(\Delta n^2) \) for each pair \( (In(v), Out(v)) \). Our total space for all the pairs \( (In(v), Out(v)) \), including the global incidence matrix, is now \( O(n^2) \).

\( L(u, w, d) \) Instead of the lists \( L(u, w, d) \) of vertices \( v \) with \( dist_v(u, w) = d \), we maintain only the oldest witness \( v \) with \( dist_v(u, w) \leq d \). This reduces the space to a constant per list, hence to to \( O(\Delta n^2) \), as desired.

The switch from `=' to `\leq' is crucial to our amortization. The point is that if \( v \) is not a witness for \( dist_v(u, w) \leq d \), it will not become so before \( (In(v), Out(v)) \) is replaced by a multi-insert.

We maintain a historical list of the vertices ordered according to dates. Note that this order has nothing to do with the order \( \prec \) used previously. When we do a multi-insert of edges around \( v \), we move \( v \) to the end of the historical list.

Suppose \( d = dist_v(u, w) \) is increased. If \( v \) was the oldest witness of \( (u, w, d) \), we scan the historical list starting from the successor of \( v \), stopping as soon as we find a younger vertex \( v' \) with \( dist_v'(u, w) = d \). Then \( v' \) is the new oldest witness of \( (u, w, d) \), and otherwise, there is no witness. If \( dist_v(u, w) \) was increased by more than one, we repeat the above scanning if \( v \) was also the oldest witness for \( (u, w, d + 1) \). As soon as we reach some value \( e \) so that \( v \) was not the oldest witness for \( (u, w, e) \), we stop; if \( v' \neq v \) is the oldest witness for \( (u, w, e) \), it is also an older witness for all \( (u, w, f) \) with \( f > e \).

Since we always scan younger and younger witnesses, we see that a vertex \( v \), between multi-inserts around \( v \), can only be scanned once for each triple \( (u, w, d) \). So, the amortized cost of each multi-insert is \( O(\Delta n^2) \). By getting rid of most pointers, we also reduce the constants hidden in the \( O \)-notation.

Below, we briefly describe how the above techniques are applied by King [9] to transitive closure, and leave the discussion of approximate and exact shortest paths to the full paper. We show that our space improvements remain valid for these applications.

### 3.2 Transitive Closure

King maintains the transitive closure of a digraph, in a manner similar to the classical method of repeated squaring of matrices. For \( i = 0, 1, \ldots, h = \lceil \log n \rceil \),
we maintain a digraph $G^n$ that is a subgraph of the transitive closure, and which is guaranteed to have an edge $(v, w)$ if there is a dipath from $v$ to $w$ in $G$ of length at most $2^i$, i.e., $G^0 = G$, and $G^h$ is the transitive closure.

For $i > 0$, and for each vertex $v \in V$, we maintain the pair $(In^{i+1}(v'), Out^{i+1}(v'))$, with $\Delta = 2$. For $i \geq 1$, the edges of $G^{i+1}$ are defined to be all pairs $(u, w)$ such that there is a path from $u$ to $w$ in $(In^i(v), Out^i(v))$ of length 3 or less.

**Lemma 2 (9).** If a dipath $P$ from $u$ to $w$ has length at most $2^{i+1}$ and the youngest vertex in the dipath is $v$, there is a path from $u$ to $w$ in $(In^{i+1}(v), Out^{i+1}(v))$ of length 3 or less.

**Proof.** Let $v'$ be a middle vertex in $P$. Assume $v'$ is in $Out^{i+1}(v)$. Then $P[u, v], P[v', v], P[w', w]$ are all of length at most $2^i$. Hence $(u, v), (v, v')$ and $(v', w)$ were all in $G^i$ when $(In^{i+1}(v), Out^{i+1}(v))$ was last reinitialized.

To save space, we implement the $(In^{i+1}(v'), Out^{i+1}(v'))$ pairs as in the previous section, with incidence matrices for $i = 1, ..., h$. When a multi-insert of edges around $v$ occurs, and $G^0$ has been updated, then for $i = 1, ..., h$, we reinitialize only $(In^{i+1}(v), Out^{i+1}(v))$ using all edges currently in $G^i$. A the multi-insert may create edges in $G^i$ not incident to $v$ which do not affect older pairs $(In^{i+1}(v'), Out^{i+1}(v'))$.

If an edge is deleted from $G$, it is deleted from the incidence matrix for $G^0$. For $i = 1, ..., h$, as edges in $G^{i−1}$ are deleted, this may result in the destruction of paths in $(In^{i+1}(v'), Out^{i+1}(v'))$ which in turn may result in the deletion of edges in $G^{i+1}$.

Given any edge $(u, w) \in G^{i+1}$, to expand it to a path, we consider the witness $v$ for $\text{dist}_i^{i+1}(u, w) \leq 3$. Then $(In^{i+1}(v), In^{i+1}(v))$ provide us a path of length $\leq 3$ from $(u, w)$ using edges from $G^i$, and each of these edges can be expanded recursively.

To get simple paths, we combine with a linear time computation of strongly connected components [12]. For each strongly connected component, using an intree and an out-tree from an arbitrary vertex, we provide a simple path between any pair of vertices in the same strongly connected component. In linear time, we also label each vertex with the strongly connected component it is in. Now we only need to reconstruct segments between strongly connected components. So, if $G^{i+1}$ provides us with a path of edges from $G^i$, we only expand edge $(x, y)$ if $x$ and $y$ are in different connected components.

Since there are $\log n$ levels, we maintain the transitive closure in $O(n^2 \log n)$ amortized time per multi-insert and $O(n^2 \log n)$ space.

4 Demetrescu-Italiano Transitive Closure

The Demetrescu-Italiano transitive closure algorithm represents the graph as an adjacency matrix and achieves a $\log n$ speed-up by basing its algorithm on the static recursive scheme [8,10] for computing transitive closure, rather than
repeated squaring. The adjacency matrix of the graph is subdivided into four sub-
matrices. The transitive closure can be shown to be a polynomial over these four 
sub-matrices. [4,3] show that the problem of updating a polynomial over Boolean 
matrices can be reduced to updating the product of degree 2 polynomials, with 
some errors allowed, as described below. Solving this problem in \( O(n^2) \) space 
gives an \( O(n^2) \) space bound for the whole algorithm.

Let \( M_1 \) and \( M_2 \) be two \( n \times n \) Boolean matrices. In computing their product, 
there are two kinds of updates allowed to either matrix in which 0’s are changed 
to 1’s: one where any entry can be set to 1 and one where just entries in a 
particular row \( i \) and column \( i \) can be set to 1. The first is done by LazySet, the 
second by SetRow\( (i,M) \) and SetCol\( (i,M) \), respectively. In addition, 1’s can be 
changed to 0’s by ReSet. The “error” allowed is that in certain circumstances, 
the bits set by LazySet can be ignored.

When entries \( (x,y) \) in \( M_1 \) and \( (y,z) \) in \( M_2 \) are both 1, we call these entries 
a witness pair for \( (x,z) \), and \( y \) a witness for \( (x,z) \). A witness pair \( (x,z) \) cannot 
be ignored if its two entries \( (x,y) \) and \( (y,z) \) exist currently and they existed 
since a time when one of them was contained in a row or column updated by 
a SetRow or SetCol. That is, the problem is to maintain for each entry in the 
matrix product a witness pair \( (x,y,z) \) if one exists and was created at or before 
the most recently executed SetRow\( (x,M_1) \), SetCol\( (y,M_1) \), SetRow\( (y,M_2) \), or 
SetCol\( (z,M_2) \).

The goal is to spend no more than \( O(n^2) \) time per SetCol and SetRow. 
Essentially, LazySet does nothing except change bits in \( M_1 \) and \( M_2 \); SetCol and 
SetRow do the work. The cost of ReSet is amortized against the operations which 
change the 0’s to 1’s.

The implementation of this in \( O(n^3) \) space is straightforward, and involves 
the keeping of lists of relevant witnesses.

Our result: We can accomplish the multiplication of two matrices (with the 
allowable errors) more easily while maintaining witnesses. We do this with no 
change to the asymptotic time, maintaining a cost of \( O(n^2) \) time per SetRow 
and SetCol, and space \( O(n^2) \). We use the technique of the previous sections, of 
keeping a pointer to the next available witness. We keep:

- For each \( x,z \), a pointer \( P1(x,z) \) which is initially set to the first column of \( y \) 
in \( M_1 \) such that \( M_1(x,y) \) AND \( M_2(y,z) \); these two entries are a witness pair for 
\( (x,z) \)
- For each \( x,z \), a pointer \( P2(x,z) \) which is initially set to the first row \( y \) in \( M_2 \) 
such that \( M_1(x,y) \) AND \( M_2(y,z) \); these are a witness pair for \( (x,z) \).
- A queue of \( y \) ’s ordered by order of last execution of either SetCol\( (y,M_1) \) or 
SetRow\( (y,M_2) \).
- For each \( (x,z) \), a pointer \( P3(x,z) \) which is initially set to the first \( y \) in the 
queue such that \( M_1(x,y) \) AND \( M_2(y,z) \)
- Back pointers from the entries which are witnesses back to the \( (x,z) \) for which 
they are a witness.

To do updates:
When SetRow\( (x) \) in \( M_1 \) is executed, we reset \( P1(x,z) \) to the beginning for all \( z \).
When SetCol($z$) in $M_2$ is executed, we reset $P_2(x, z)$ to the beginning for all $x$. When $M_1(x, y)$ or $M_2(y, z)$ is set to 0 during a ReSet:

- If $P_1(x, z)$ points to $y$, then we move $P_1(x, z)$ forward until another witness is found or the end of the row.
- If $P_2(x, z)$ points to $y$, then we move $P_2(x, z)$ forward until another witness is found or the end of the column.
- If $P_3(x, z)$ points to $y$, then we advance $P_3(x, z)$ until another witness is found.

Running time analysis: Each time a pointer is set back to the beginning of a row or column, there may be another $n$ checks at a cost of $O(1)$ each. $P_3(x, z)$ is never set back but there may be an added cost of $O(1)$ for every pair $(x, z)$ for each queue change. We charge the SetCol or SetRow to cover this.

References

Finding the Most Vital Node of a Shortest Path

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Abstract. In an undirected, 2-node connected graph $G = (V, E)$ with positive real edge lengths, the distance between any two nodes $r$ and $s$ is the length of a shortest path between $r$ and $s$ in $G$. The removal of a node and its incident edges from $G$ may increase the distance from $r$ to $s$. A most vital node of a given shortest path from $r$ to $s$ is a node (other than $r$ and $s$) whose removal from $G$ results in the largest increase of the distance from $r$ to $s$. In the past, the problem of finding a most vital node of a given shortest path has been studied because of its implications in network management, where it is important to know in advance which component failure will affect network efficiency the most. In this paper, we show that this problem can be solved in $O(m + n \log n)$ time and $O(m)$ space, where $m$ and $n$ denote the number of edges and the number of nodes in $G$.

1 Introduction

The computational infrastructure throughout society is becoming increasingly large and complex. Networks of workstations are vulnerable to attack and failure, and it is generally recognized that the survivability of our computing systems is a critical issue (IEEE Computer dedicated its August 2000 issue to this topic). We are interested in a particular type of survivability: How is a communication network affected by the failure of a component? In this paper, we consider the effect that a node failure will have on a shortest path between two nodes. Our scenario assumes that each message is routed along a shortest path in a communication graph from its source to its destination. When a node on that path fails, we need to replace the old route by a new one, preferably by a shortest path in the graph that does not contain the failed node. Let us call this new route a replacement (shortest) path; it will in general be longer than the path.

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Finding the Most Vital Node of a Shortest Path

it replaces, but it certainly will never be shorter. From a network management point of view, it is desirable to know for a shortest path ahead of time which node failure will result in the longest replacement path. Such a node is called a *most vital node*, because its failure degrades the transmission from the source to the destination most strongly.

The problem of finding a most vital node of a shortest path has been defined and motivated by Corley and Sha [3]. More precisely, they considered the more general problem of finding the \( k \) most vital nodes of a shortest path, that is the \( k \) nodes whose removal will increase the distance between the source and the destination node the most, and they gave an exponential algorithm for solving the problem. For the case \( k = 1 \), an efficient implementation of their algorithm requires \( O(mn + n^2 \log n) \) time, where \( m \) and \( n \) denote the number of edges and the number of nodes in the underlying graph. Note that this is not better than the trivial bound that we get by recomputing from scratch the replacement shortest path for every node along the given shortest path. Later on, Bar-Noy et al. [2] proved that, for arbitrary \( k \), the problem is strongly NP-hard. Finally, Venema et al. [13] studied the problem for \( k = 1 \) in a parallel computing environment, providing a polynomial algorithm.

In a related scenario, nodes are reliable, but edges can fail. The problem of finding a *most vital edge* on a shortest path has been studied extensively in the past: We look for an edge whose failure leads to the longest replacement path [1,2,3]. Now, naturally a replacement path is just a path avoiding the failed edge. Let us assume that the source and destination nodes lie in a 2-edge connected component of the given graph; otherwise, a most vital edge is just a bridge between these nodes, and that is easy to find. The fastest algorithm to compute most vital edges on a pointer machine runs in \( O(m + n \log n) \) time and \( O(m) \) space [6]. On a RAM, there is an algorithm to solve the problem (and an interesting variant of it defined in [7]) in \( O(m \cdot \alpha(m, n)) \) time and \( O(m) \) space [8], where \( \alpha(m, n) \) denotes the functional inverse of the Ackermann function defined in [12].

In this paper, we show that the problem of finding a most vital node for a shortest path in a 2-node connected, undirected and positively weighted graph can be solved on a pointer machine in \( O(m + n \log n) \) time and \( O(m) \) space. The efficiency of our algorithm is based on two considerations. First, we make use of specific structural properties of replacement paths in the computation. This is realized by means of a priority queue that stores certain distance values for certain nodes. Unfortunately, the priority queue contains distance values that would lead to an incorrect result if they ever would be used, and it also contains the desired distance values leading to the correct result. The reason for this mix is algorithmic performance: We have no way of efficiently distinguishing between both, but we make sure that the algorithm never uses the undesired values. Second, we perform this computation incrementally as we visit the nodes along the shortest path.

The paper is organized as follows: In Section 2 we define the problem formally and give the required basic definitions. In Section 3, we present the structural
properties of replacement paths, along with our algorithm. Finally, Section 4 discusses modifications and further applications, and lists some open problems.

2 Basic Definitions

Let \( G = (V, E) \) be an undirected graph, where \( V \) is the set of nodes and \( E \subseteq V \times V \) is the set of edges. Let \( n \) and \( m \) denote the number of nodes and the number of edges, respectively, and, for each \( e \in E \), let \( w(e) \) be a positive real length. A graph \( H = (V(H), E(H)) \) is called a subgraph of \( G \) if \( V(H) \subseteq V \) and \( E(H) \subseteq E \). If \( V(H) = V \) then \( H \) is called a spanning subgraph of \( G \).

A simple path (or a path for short) in \( G \) is a subgraph \( P \) with \( V(P) = \{v_0, v_1, \ldots, v_k | v_i \neq v_j \text{ for } i \neq j \} \) and \( E(P) = \{(v_i, v_{i+1}) | 0 \leq i < k \} \), also denoted as \( P(v_0, v_k) = \{v_0, \ldots, v_k\} \). Path \( P(v_0, v_k) \) is said to go from \( v_0 \) to \( v_k \) or, alternatively, to be between \( v_0 \) and \( v_k \). Its length is the sum of the lengths of the path edges, and will be denoted as \( |P(v_0, v_k)| \). A graph \( G \) is connected if, for any two nodes \( u, v \in V \), there exists a path in \( G \) going from \( u \) to \( v \). A connected acyclic spanning subgraph of \( G \) is called a spanning tree of \( G \). Let \( G - v \) denote the graph obtained by removing from \( G \) the node \( v \) and its incident edges. A graph \( G \) is 2-node connected if for any \( v \in V \), \( G - v \) is connected.

A path between two nodes \( r \) and \( s \) is shortest in \( G \) if it has minimum length among all the paths in \( G \) between \( r \) and \( s \). In this case, we denote the path by \( P_G(r, s) \), and its length, also known as the distance in \( G \) between \( r \) and \( s \), by \( d_G(r, s) \). For a distinguished node \( r \in V \), called the source, and all the nodes \( v \in V \setminus \{r\} \), a single-source shortest paths tree (SPT) \( S_G(r) \) in \( G \) is a spanning tree of \( G \) rooted in \( r \) and formed by the union of shortest paths, with one shortest path from \( r \) to \( v \) for each \( v \in V \setminus \{r\} \).

Let \( P_{G-v}(r, s) \) be a shortest path between \( r \) and \( s \) in \( G - v \), named a replacement shortest path for \( v \), and let \( d_{G-v}(r, s) \) denote its length. The most vital node (MVN) problem with respect to \( P_G(r, s) \) asks for finding a node \( v^* \in V \setminus \{r, s\} \) such that \( d_{G-v^*}(r, s) \geq d_{G-v}(r, s) \), for any \( v \in V \setminus \{r, s\} \).

3 An Efficient Solution of the MVN Problem

Let \( P_G(r, s) = \langle v_0, v_1, \ldots, v_k \rangle \) be a shortest path between \( r = v_0 \) and \( s = v_k \) in \( G \). First of all, notice that a node (other than \( r \) and \( s \)) whose removal increases the distance between \( r \) and \( s \) must belong to the node set \( \{v_1, \ldots, v_{k-1}\} \). Therefore, in the following, we will consider only the removal of the nodes along the path.

3.1 The Structural Properties of Replacement Shortest Paths

In this section, we present the structural properties of replacement shortest paths that will form the basis for the efficiency of our algorithm.

Let \( S_G(r) \) denote a SPT in \( G \) rooted at \( r \) and containing \( P_G(r, s) \), and let \( v_i, 1 \leq i \leq k - 1 \), be a node on \( P_G(r, s) \). When node \( v_i \) and its incident edges are removed from \( G \), \( S_G(r) \) is partitioned into a set of subtrees, that we classify as follows (see Figure 1):
Fig. 1. Node \( v_i \) is removed from \( G \): \( S_G(r) \) is partitioned into a set of subtrees, with node sets \( U_i, O_i \) and \( D_i \).

1. the subtree of \( S_G(r) \) containing the parent \( v_{i-1} \) of \( v_i \); we call the nodes of this subtree the \textit{upwards nodes} of \( v_i \), and we denote them as \( U_i \);
2. the subtree of \( S_G(r) \) containing the child \( v_{i+1} \) of \( v_i \); we call the nodes of this subtree the \textit{downwards nodes} of \( v_i \), and we denote them as \( D_i \);
3. all the remaining subtrees of \( S_G(r) \); we call the nodes of the union of all these subtrees the \textit{outwards nodes} of \( v_i \), and we denote them as \( O_i \).

In the rest of the paper, we will make use of the following properties, that hold for \( i, j = 1, \ldots, k - 1 \) and \( j \neq i \):

(P1:) \( U_i \cup O_i \cup D_i = V \setminus \{v_i\} \);
(P2:) \( U_i, O_i \) and \( D_i \) are pairwise disjoint;
(P3:) \( U_i \subseteq U_{i+1} \);
(P4:) \( D_{i+1} \subseteq D_i \);
(P5:) \( O_i \cap O_j = \emptyset \).

We start by observing that for nodes in \( U_i \), the shortest path to \( r \) does not contain \( v_i \). This immediately implies the following:

**Lemma 1.** For each node \( u \in U_i \), \( d_G(r, u) = d_{G-v_i}(r, u) \).

On the other hand, for nodes in \( D_i \), we have that the distance to \( s \) does not change when \( v_i \) is removed:

**Lemma 2.** For each node \( u \in D_i \), \( d_G(s, u) = d_{G-v_i}(s, u) \).

**Proof.** Suppose, for the sake of contradiction, that \( d_G(s, u) \neq d_{G-v_i}(s, u) \). Let \( S_G(s) \) be a SPT in \( G \) rooted in \( s \) containing \( P_G(r, s) \). From \( d_G(s, u) \neq d_{G-v_i}(s, u) \),
Lemma 3. Any replacement shortest path for node $v_i$ can be expressed as a concatenation of $P_{G-v_i-D_i}(r,x)$, edge $(x,y)$ and $P_G(y,s)$, where $y$ is the entry node into $D_i$, $P_{G-v_i-D_i}(r,x)$ is a shortest path from $r$ to $x$ in $G - v_i - D_i$, and $P_G(y,s)$ is a shortest path from $y$ to $s$ in $G$.

Proof. Since $r$ is not contained in $D_i$, but $s$ is, there is a first node on the path traced from $r$ towards $s$ that belongs to $D_i$. Call that node $y$, and call its predecessor on the path $x$. This proves the first and second part of the claim. Part three is due to Lemma 2.

\section{3.2 Computing Components of Distances}

Lemma 3 allows us to compute replacement shortest paths as follows. First, we compute in $G$ a SPT rooted in $s$. This gives us all distances $d_{G-v_i}(y,s)$ for
For \( x \in \mathcal{U}_i \), from Lemma 1 we have \( d_{G-v_i-D_i}(r,x) = d_G(r,x) \), and therefore the SPT rooted in \( r \) that contains \( P_G(r,s) \), denoted as \( S_G(r) \), gives us all these values. The remaining more interesting task is the computation of \( d_{G-v_i-D_i}(r,x) \) for \( x \in \mathcal{O}_i \). We propose to do this as follows, making use of \( S_G(r) \). When node \( v_i, 1 \leq i \leq k-1 \), is removed, we consider the subtree of \( S_G(r) \) induced by \( \mathcal{U}_i \) – which is of course a SPT rooted in \( r \) of the subgraph of \( G \) induced by the node set \( \mathcal{U}_i \). Then we compute the distance from \( r \) to all the nodes in \( \mathcal{O}_i \) in the subgraph of \( G \) induced by \( \mathcal{U}_i \cup \mathcal{O}_i \). We do this by applying Dijkstra’s algorithm \([4]\) in the following way to the nodes in \( \mathcal{O}_i \), starting from the precomputed distances for \( \mathcal{U}_i \). Let \( E(\mathcal{U}_i, \mathcal{O}_i) \) be the subset of edges in \( E \) having one end node in \( \mathcal{U}_i \) and the other one in \( \mathcal{O}_i \), let \( E(\mathcal{U}_i, x) \) be the subset of edges in \( E(\mathcal{U}_i, \mathcal{O}_i) \) having one end node in \( \mathcal{U}_i \) and the other one in \( x \in \mathcal{O}_i \), and let \( E(\mathcal{O}_i, \mathcal{O}_i) \) be the subset of edges in \( E \) having both end nodes in \( \mathcal{O}_i \). We create an initially empty heap \( \mathcal{H}_i \), inserting into it all the nodes \( x \in \mathcal{O}_i \), with key

\[
    k(x) = \begin{cases} 
        \min_{f = (u,x) \in E(\mathcal{U}_i,x)} \{d_{G-v_i}(r,u) + w(f)\} & \text{if } E(\mathcal{U}_i,x) \neq \emptyset; \\
        +\infty & \text{otherwise}. 
    \end{cases} 
\]

Afterwards, we extract the minimum \( k(x) \) from \( \mathcal{H}_i \); this gives us \( d_{G-v_i-D_i}(r,x) \). Then, we update the keys of adjacent nodes still appearing in \( \mathcal{H}_i \), by making use of edges in \( E(\mathcal{O}_i, \mathcal{O}_i) \), exactly as in Dijkstra’s algorithm. The algorithm iterates until \( \mathcal{H}_i \) is empty.

This algorithm has an efficient implementation, as expressed in the following lemma:

**Lemma 4.** The values \( d_{G-v_i-D_i}(r,x) \) for all nodes \( x \in \mathcal{O}_i, i = 1, \ldots, k-1 \), can be computed in \( O(m + n \log n) \) time and \( O(m) \) space.

**Proof.** The initial computation of \( S_G(r) \) takes \( O(m + n \log n) \) time and \( O(m) \) space \([5]\). Throughout the \( k-1 \) iterations in our algorithm, \( k-1 = O(n) \) heaps are created. Let \( n_i \) denote the number of nodes of \( \mathcal{O}_i \), and let \( m_i = |E(\mathcal{U}_i, \mathcal{O}_i) \cup E(\mathcal{O}_i, \mathcal{O}_i)| \). On heap \( \mathcal{H}_i \), we perform \( O(n_i) \) Insert operations, and from Lemma 1 key initialization can be performed in \( O(m_i) \) time once \( S_G(r) \) has been computed. Moreover, we perform \( O(n_i) \) ExtractMin and \( O(m_i) \) DecreaseKey operations. By using Fibonacci heaps \([5]\), for all the nodes \( x \in \mathcal{O}_i \), \( d_{G-v_i-D_i}(r,x) \) can then be computed in \( O(m_i + n_i \log n) \) time.

Since each DecreaseKey operation is associated with an edge of \( G \), and each edge of \( G \) is considered at most twice, and given that sets \( \mathcal{O}_i \) are disjoint, we finally have that the total time is

\[
    \sum_{i=1}^{k-1} O(m_i + n_i \log n_i) = O(m + n \log n). \quad \square
\]
3.3 Combining Components of Distances

We are now ready to combine the distance components computed so far. We first give a description of the algorithm, and we then analyze its correctness and its time and space complexity.

We consider the nodes \( v_1, \ldots, v_{k-1} \) in this order along \( P_G(r, s) \), and when the node \( v_i \) is considered, we maintain in a heap \( H \) the set of nodes \( D_i \) associated with it. For each node \( y \in D_i \), we consider the subset of edges in \( E(U_i \cup O_i, D_i) \) incident to \( y \), denoted as \( E(U_i \cup O_i, y) \). In the heap, with node \( y \) a key \( k(y) \) is associated that satisfies the following condition immediately before a \textit{FindMin} operation on \( H \) is performed:

\[
k(y) = \min_{f = (x, y) \in E(U_i \cup O_i, y)} \{d_{G-v_i-D_i}(r, x) + w(f) + d_G(y, s)\}. \tag{3}
\]

Notice that in general, this key value is not the length of a shortest path in \( G-v_i \) from \( r \) to \( s \) through \( y \), but, as we explained in Section 3.1, we cannot afford to maintain these latter values. We will show later that these keys, however, give us sufficient information to solve the problem.

The algorithm works in stages. At the beginning, the heap \( H \) is created for \( D_0 \), that is, all the nodes in the subtree \( D_0 \) rooted at \( v_1 \) in \( S_G(r) \) are inserted, with arbitrarily large keys associated. At the \( i \)th stage, we consider the node \( v_i \) on \( P_G(r, s) \), and we update the heap in the following way:

**Step 1:** We remove from \( H \) the node \( v_i \) and the nodes \( O_i \) associated with it. (Comment: This leaves exactly the nodes in \( D_i \) in \( H \); we update their keys in Steps 2 and 3.)

**Step 2:** We consider all the nodes in \( O_i \); for each such node \( x \), we inspect its incident edges, and we limit further actions on those crossing into \( D_i \). Let \( f = (x, y) \) be one of these crossing edges, if any, and let

\[
k' = d_{G-v_i-D_i}(r, x) + w(f) + d_G(y, s), \tag{4}
\]

where \( d_{G-v_i-D_i}(r, x) \) has been computed by means of the procedure described in Section 3.2. If \( k' < k(y) \), we decrease the key of \( y \) in \( H \) to value \( k' \). (Comment: When this step is completed, all the crossing edges associated with \( v_i \) and induced by its removal have been exhausted.)

**Step 3:** We then consider all the nodes in \( v_{i-1} \cup O_{i-1} \); for each node \( x \) in this set, we look at its incident edges, and we limit further actions on those crossing into \( D_i \). Let \( f = (x, y) \) be one of these crossing edges, if any, and let

\[
k' = d_G(r, x) + w(f) + d_G(y, s). \tag{5}
\]

If \( k' < k(y) \), we decrease the key of \( y \) in \( H \) to value \( k' \). (Comment: When this step is completed, all the crossing edges associated with \( v_i \) and induced by the reinsertion of \( v_{i-1} \) have been exhausted, and the corresponding key maintenance in the heap is complete.)
Step 4: We finally find the minimum of $H$. (Comment: We will prove shortly that the key associated with this minimum is exactly the length of a replacement shortest path in $G - v_i$ between $r$ and $s$, that is $d_{G - v_i}(r, s)$.)

When all stages $1, \ldots, k - 1$ have been completed, a most vital node can then be determined as a node $v_j$ on $P_G(r, s)$ such that

$$d_{G - v_j}(r, s) = \max_{i=1, \ldots, k-1} \{d_{G - v_i}(r, s)\}. \quad (6)$$

Let us now prove that our algorithm indeed computes at each stage the length of a corresponding replacement shortest path.

Lemma 5. The minimum key found in $H$ at the $i$th stage is the length of a replacement shortest path between $r$ and $s$ in $G - v_i$.

Proof. We prove the lemma in two steps. First, we prove that each key in the heap $H$, say $k(y)$ for node $y$, is the length of a shortest path in $G - v_i$ from $r$ to $s$ through the entry node $y$. The reason is that our algorithm inspects all crossing edges $(x, y)$ incident to $y$, and keeps track of the best.

Second, we prove that at least one node in the heap has a key corresponding to the length of a replacement shortest path between $r$ and $s$ in $G - v_i$. In fact, for any replacement shortest path $P$ in $G - v_i$, the corresponding entry node $y$ is in $H$. Let $x$ be its predecessor on $P$. Then, $k(y)$ equals the length of $P$, because the prefix of $P$ from $r$ to $x$ is contained in $G - v_i - D_i$, and then $d_{G - v_i - D_i}(r, x) = d_{G - v_i}(r, x)$. Therefore, (11) and (3) are both minimized when edge $(x, y)$ is considered, and $k(y) = d_{G - v_i}(r, s)$. \qed

The following theorem can finally be proved:

Theorem 1. A most vital node on a shortest path $P_G(r, s)$ between two nodes $r$ and $s$ in a 2-node connected, undirected graph $G = (V, E)$ with $n$ vertices and $m$ edges, with positive real edge lengths, can be determined in $O(m + n \log n)$ time and $O(m)$ space.

Proof. The correctness of the above algorithm derives from Lemma 5. The time complexity follows from that of Lemma 4 for the initial phase. This allows us to compute (4) in $O(1)$ time for each crossing edge. Clearly, (5) can be computed in $O(1)$ time for each crossing edge as well, once $S_G(r)$ and $S_G(s)$ have been computed. Globally, we perform $O(m)$ computations of (11) and (5), since a crossing edge is checked at most twice, once each in Step 2 and Step 3. Then, we make use of a Fibonacci heap [5] for maintaining $H$. Since each node of $G$ is inserted into the heap and removed from it at most once, we have a single Make-Heap, $O(n)$ Insert, $k - 1 = O(n)$ FindMin, $O(n)$ Delete and $O(m)$ DecreaseKey operations (since a key may be decreased only when a new crossing edge is considered), and thus we obtain a total time of $O(m + n \log n)$ for heap operations. The time complexity for other tasks is respectively $O(m + n \log n)$ time for computing $S_G(r)$ and $S_G(s)$, $O(n)$ time for managing sets $O_i$, $i = 1, \ldots, k - 1$, and $O(n)$ time for computing (6). Finally, $O(m)$ space is trivially enough to handle all the operations. Thus, the claim follows. \qed
4 Discussion

In this paper we have presented a fast solution to the problem of finding a most vital node along a shortest path $P_G(r, s)$ between two nodes $r$ and $s$ in a graph $G$. Implicitly, our algorithm computes not only the lengths of replacement paths, but also the paths themselves; it can be easily modified to do so explicitly. It runs in $O(m + n \log n)$ time and $O(m)$ space, which, as far as we know, is the first improvement over the trivial bound of $O(nm + n^2 \log n)$ time and $O(m)$ space that we get by recomputing a replacement shortest path between $r$ and $s$ from scratch after the removal of each node along $P_G(r, s)$.

In some applications, such as transportation networks, it appears to be more realistic to associate costs with both, nodes and edges, instead of only one type of network components. Our approach also answers the corresponding more general question that suggests itself: In a graph where both edges and nodes have a non-negative cost, and where both edges and nodes can fail, what is a most vital edge or node on a shortest path? The algorithm can be modified slightly and still runs within the same asymptotic bounds for this more general question, for two reasons. First, edge failures can be modelled as node failures, when each edge is replaced by a path of length two with an extra node in the center of that path; then, the failure of the extra node represents the failure of the original edge. Second, Dijkstra’s algorithm can be adapted easily to work also for shortest path computations in graphs with costs on edges and nodes, where the cost of a path is the sum of the costs of its edges and nodes. Obviously, both modifications do not change the asymptotic bounds for the runtime and the storage space.

Our algorithmic solution is also useful in quite a different application context. In large networks, components (nodes and edges) may be owned by different owners. The incentive of an owner of a component to forward a message, naturally, is to get some reward. In standard economic terms, that reward is the price of the service of forwarding the message. It is economically desirable that each owner declares the true price for the service that its component offers, so as to allocate the overall resources in a best possible way. Nevertheless, there is an incentive for owners to speculate and ask for a higher price, in the hope of getting a higher profit. This leads to economically suboptimal resource allocation and is therefore undesirable. A few studies in the computer science literature have devoted their attention to setting the boundary conditions in such a way that speculating with high prices does not pay off. This is known as mechanism design for selfish agents [9,10,11]. In [10], Nisan is explicitly suggesting a rewarding model for forwarding messages on paths, based on microeconomic theory, that requires the computation of replacement path lengths for edges. This model assumes that only edges charge a price for forwarding a message; nodes perform their service for free. Here, again, it would be more realistic to have a price for both, edges and nodes, than a limitation of the pricing to the edges alone. A straightforward modification of the charging scheme from [10] to node and edge prices serves the desired purpose. Now, the modification of our algorithm for node and edge costs and failures is an efficient implementation of the required replacement path computations.
Our solution is efficient, but it is still open whether it is optimal. Notice that to improve our solution, a faster computation of a single source shortest paths tree must be provided. For more general settings there are still many open problems; one of them deals with multiple edge or node failures on a shortest path.

References

Algorithm for the Cost Edge-Coloring of Trees

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Abstract. Let $C$ be a set of colors, and let $\omega$ be a cost function which assigns a real number $\omega(c)$ to each color $c$ in $C$. An edge-coloring of a graph $G$ is to color all the edges of $G$ so that any two adjacent edges are colored with different colors. In this paper we give an efficient algorithm to find an optimal edge-coloring of a given tree $T$, that is, an edge-coloring $f$ of $T$ such that the sum of costs $\omega(f(e))$ of colors $f(e)$ assigned to all edges $e$ is minimum among all edge-colorings of $T$. The algorithm takes time $O(n\Delta^2)$ if $n$ is the number of vertices and $\Delta$ is the maximum degree of $T$.

1 Introduction

A vertex-coloring of a graph $G = (V,E)$ is to color all the vertices of $G$ so that any two adjacent vertices are colored with different colors. We denote by $n$ the number of vertices in $G$. Let $C = \{c_1, c_2, \ldots, c_m\}$ be a set of colors, and assume that $m = |C|$ is sufficiently large, say $m = n$. Let $\omega : C \rightarrow \mathbb{R}$ be a cost function which assigns a real number $\omega(c) \in \mathbb{R}$ to each color $c \in C$. The cost vertex-coloring problem is to find a vertex-coloring $f : V \rightarrow C$ of $G$ such that $\sum_{v \in V} \omega(f(v))$ is as small as possible. One can observe that the cost vertex-coloring problem is NP-hard since the ordinary vertex-coloring problem is NP-hard. The cost vertex-coloring problem remains NP-hard for interval graphs, but the problem can be solved in linear time for trees [9]. The “vertex-chromatic sum problem” in [10,11] is merely the cost vertex-coloring problem for a special case where $\omega(c_i) = i$ for each index $i \geq 1$. Jansen gives a polynomial-time algorithm to solve the vertex-chromatic sum problem for graphs of bounded treewidth [5].

In this paper we introduce the cost edge-coloring problem. An edge-coloring $f : E \rightarrow C$ of a graph $G = (V,E)$ is to color all the edges of $G$ so that any two adjacent edges are colored with different colors. The minimum number of colors required for edge-colorings of $G$ is called the chromatic index, and is denoted by $\chi'(G)$. An ordinary edge-coloring problem is to compute the chromatic index $\chi'(G)$ of a given graph $G$ and find an edge-coloring of $G$ with $\chi'(G)$ colors. On the other hand, the cost edge-coloring problem is to find an optimal edge-coloring of $G$, that is, an edge-coloring $f$ such that $\sum_{e \in E} \omega(f(e))$ is minimum among all edge-colorings of $G$. An optimal edge-coloring does not always use the minimum number $\chi'(G)$ of colors. For example, suppose that $\omega(c_1) = 1$ and $\omega(c_i) = 5$ for each index $i \geq 2$, then the graph $G$ with $\chi'(G) = 3$ in Figure 1(a) can be
Algorithm for the Cost Edge-Coloring of Trees 289

Fig. 1. (a) Edge-coloring with $\chi'(G)$ colors and (b) optimal edge-coloring with $\chi'(G) + 1$ colors.

uniquely colored with the three cheapest colors $c_1, c_2$ and $c_3$ as in Figure 1(a), but the edge-coloring is not optimal; an optimal edge-coloring of $G$ uses the four cheapest colors $c_1, c_2, c_3$ and $c_4$ as depicted in Figure 1(b). The edge-chromatic sum problem introduced by Giaro and Kubale [4] is merely the cost edge-coloring problem for a special case where $\omega(c_i) = i$ for each index $i \geq 1$. The cost edge-coloring problem has a natural application in scheduling theory. Consider the scheduling of unit execution time of biprocessor tasks on dedicated machines. An example of such tasks is the file transfer problem in a computer network in which each file engages two corresponding nodes, sender and receiver, simultaneously [3]. Another example is the biprocessor diagnostic problem in which links execute concurrently the same test for a fault tolerant multiprocessor system [6]. These problems can be modeled by a graph $G$ in which machines correspond to the vertices and tasks correspond to the edges. An edge-coloring of $G$ corresponds to a schedule, where the edges colored with color $c_i \in C$ represent the collection of tasks that are executed in the $i$th time slot. For each $i$, if a task is executed in the $i$th time slot, then it takes the cost $\omega(c_i)$. The goal is to find a schedule that minimizes the total cost. This corresponds to the cost edge-coloring problem.

One can observe that the cost edge-coloring problem is NP-hard since the ordinary edge-coloring problem is NP-hard. Furthermore, the cost edge-coloring problem remains NP-hard for bipartite graph, because the edge-chromatic sum problem is NP-hard for bipartite graphs [4]. When restricted to trees, many NP-hard problems can be efficiently solved in polynomial time [1] [2] [3] [4]. Giaro and Kubale recently give an algorithm to solve the edge-chromatic sum problem for trees $T$ in time $O(n\Delta^{3.5} \log n)$, where $n$ is the number of vertices and $\Delta$ is the maximum degree of $T$ [4]. However, it is not known whether the cost edge-coloring problem can be solved in polynomial time for trees.

In this paper we give an algorithm to solve the cost edge-coloring problem for trees $T$ in time $O(n\Delta^2)$. The algorithm hence runs in linear time if $\Delta$ is a fixed constant. The algorithm takes time $O(n\Delta^{1.5} \log(\Delta N))$ if all the color costs $\omega(c)$ are integers in the range $[-N, N]$. The edge-chromatic sum problem is a special case of the cost edge-coloring problem in which $N = n$. Our algorithm can therefore solve the edge-chromatic sum problem for trees
in time $O(n \Delta^{1.5} \log n)$. The time-complexity $O(n \Delta^{1.5} \log n)$ is better than the time-complexity $O(n \Delta^{3.5} \log n)$ of Giaro and Kubale’s algorithm \[4\]. We also show that any tree $T$ has an optimal edge-coloring using exactly $\chi'(T) = \Delta(T)$ colors.

2 Preliminaries

In this section we define some terms and present easy observations. Let $T = (V, E)$ denote a tree with vertex set $V$ and edge set $E$. We will use notions such as root, child, descendant, and leaf in their usual meaning. We denote by $n$ the number of vertices in $T$. $T$ is a “free tree,” but we choose for the sake of convenience an arbitrary vertex as a root $r$, and regard $T$ as a “rooted tree.” The degree $\deg(v)$ of vertex $v$ is the number of edges incident to $v$. We denote the maximum degree of $T$ by $\Delta(T)$ or simply by $\Delta$. We denote by $d(v)$ the number of children of $v$ in $T$. Then $d(v) = \deg(v)$ if $v = r$, and $d(v) = \deg(v) - 1$ otherwise. An edge joining vertices $u$ and $v$ is denoted by $(u, v)$. For a vertex $v$ of $T$, the subtree of $T$ induced by $v$ and all descendants of $v$ is called the subtree of $T$ rooted at $v$, and is denoted by $T_v$. Clearly, $T = T_r$ for the root $r$.

Since tree $T$ has $n-1$ edges, one may assume without loss of generality that the color set $C$ satisfies $|C| = n - 1$. We however assume for the notational sake of convenience that $|C| = n$, and write $C = \{c_1, c_2, \cdots, c_n\}$. An edge-coloring $f : E \rightarrow C$ of a tree $T$ is to color all edges of $T$ by colors in $C$ so that any two adjacent edges are colored with different colors. Let $\omega : C \rightarrow \mathbb{R}$, where $\mathbb{R}$ is the set of real numbers. One may assume with loss of generality that $\omega$ is non-decreasing, that is, $\omega(c_i) \leq \omega(c_{i+1})$ for any index $i$, $1 \leq i < n$. The cost $\omega(f)$ of an edge-coloring $f$ of a tree $T = (V, E)$ is defined as follows:

$$\omega(f) = \sum_{e \in E} \omega(f(e)).$$

An edge-coloring $f$ of $T$ is called an optimal one if $\omega(f)$ is minimum among all edge-colorings of $T$. The cost edge-coloring problem is to find an optimal edge-coloring of a given tree $T$. The cost of an optimal edge-coloring of tree $T$ is called the minimum cost of tree $T$, and is denoted by $\omega(T)$.

Let $f$ be an edge-coloring of $T$. For each vertex $v$ of $T$, let $C(f, v)$ be the set of all colors that are assigned to edges incident to $v$, that is,

$$C(f, v) = \{f(e) \mid e \text{ is an edge incident to } v \text{ in } T\}.$$

We say that a color $c \in C$ is missing at $v$ if $c \notin C(f, v)$. Let Miss($f, v$) be the set of all colors missing at $v$, that is,

$$\text{Miss}(f, v) = C - C(f, v).$$

Interchanging colors in an alternating path is one of the standard techniques of an ordinary edge-coloring, which we also use in the paper. Let $f$ be an edge-coloring of a tree $T = (V, E)$, let $c_i$ and $c_j$ be any two colors in $C$, and let
Theorem 1. \( T(c_i, c_j) \) be the subgraph of \( T \) induced by all the edges colored with \( c_i \) and \( c_j \). Since \( T \) is a tree, each connected component of \( T(c_i, c_j) \) is a path of length one or more, whose edges are colored alternately with \( c_i \) and \( c_j \). We call such a path a \( c_i, c_j \)-alternating path. A vertex \( v \in V \) is an end of a \( c_i, c_j \)-alternating path if and only if exactly one of \( c_i \) and \( c_j \) is missing at \( v \). If \( v \) is an end of a \( c_i, c_j \)-alternating path, then the path is denoted by \( P(v; c_i, c_j) \). The set of all edges in \( P(v; c_i, c_j) \) is sometimes simply denoted by \( P(v; c_i, c_j) \). Interchanging colors \( c_i \) and \( c_j \) in \( P(v; c_i, c_j) \), one can obtain another edge-coloring \( f' \) of \( T \). We then have the following lemma on \( f' \).

Lemma 1. If \( c_i \in \text{Miss}(f, v), c_j \in C(f, v), \) and \( i < j \), then \( \omega(f') \leq \omega(f) \), \( c_i \in C(f', v), c_j \in \text{Miss}(f', v) \), and \( C(f', v) \cup \{c_j\} = C(f, v) \cup \{c_i\} \).

Proof. The colors of all edges assigned by \( f \) are the same as those by \( f' \) except for the edges in \( P(v; c_i, c_j) \). The first edge in \( P(v; c_i, c_j) \) is colored with \( c_j \) by \( f \), but it is colored with \( c_i \)' by \( f' \). We therefore have \( c_i \in C(f', v), c_j \in \text{Miss}(f', v) \), and \( C(f', v) \cup \{c_j\} = C(f, v) \cup \{c_i\} \). Since \( \omega(c_i) \leq \omega(c_j) \) and \( P(v; c_i, c_j) \) starts with an edge colored with \( c_j \), we have

\[
\sum_{e \in P(v; c_i, c_j)} \omega(f'(e)) \leq \sum_{e \in P(v; c_i, c_j)} \omega(f(e)).
\]

We thus have \( \omega(f') \leq \omega(f) \). \( \text{Q.E.D.} \)

3 Algorithm

In this section we prove the following theorem.

Theorem 1. An optimal edge-coloring of a tree \( T \) can be found in time \( O(n\Delta^2) \) if \( T \) has \( n \) vertices and \( \Delta \) is the maximum degree of \( T \).

In the remainder of this section we prove Theorem 1. In Section 3.1 we give an algorithm to compute the minimum cost \( \omega(T) \) of a given tree \( T \), and in Section 3.2 we give an algorithm to find an optimal edge-coloring of \( T \).

3.1 Computing the Minimum Cost \( \omega(T) \)

A “dynamic programming method” is a standard one to solve a combinatorial problem on trees. We also use it, and compute the minimum cost \( \omega(T) \) of a given tree \( T \) by the “bottom-up tree computation.” However, one can observe that the minimum cost \( \omega(T_v) \) of a subtree \( T_v \) rooted at a vertex \( v \) cannot be computed directly from the minimum costs \( \omega(T_{v_j}) \), \( 1 \leq j \leq d(v) \), of all the children \( v_j \) of \( v \). (See Figure 2)

Our first idea is to introduce a new parameter \( \omega(T_v, i) \) defined for each vertex \( v \) of \( T \) and each color \( c_i \in C \) as follows:

\[
\omega(T_v, i) = \min\{\omega(f) \mid f \text{ is an edge-coloring of } T_v \text{ and } c_i \in \text{Miss}(f, v)\}.
\]
Fig. 2. Subtree $T_v$ rooted at $v$.

Since $T$ has exactly $n-1$ edges and $|C| = n$, $T_v$ has an edge-coloring $f$ in which $c_i$ is not used by $f$ and hence $c_i \in \text{Miss}(f,v)$. Thus $\omega(T_v,i)$ is well-defined. Clearly,

$$\omega(T_v) = \min_{1 \leq i \leq n} \omega(T_v,i).$$

We compute the values $\omega(T_v,i)$ for all indices $i$, $1 \leq i \leq n$, from leaves to root $r$. Thus the DP table for each vertex $v$ consists of the $n$ values $\omega(T_v,i)$, $1 \leq i \leq n$, and has a size $n$.

The second idea is to reduce the size from $n$ to $d(v) + 1$. We first have the following two lemmas on the properties of $\omega(T_v,i)$.

**Lemma 2.** For any vertex $v$ of tree $T$, $\omega(T_v,i)$ is non-increasing with $i$, that is, $\omega(T_v,i) \geq \omega(T_v, i+1)$ for each index $i$, $1 \leq i \leq n-1$.

**Proof.** Let $f$ be an edge-coloring of $T_v$ such that $c_i \in \text{Miss}(f,v)$ and $\omega(f) = \omega(T_v,i)$. If $c_{i+1} \in \text{Miss}(f,v)$, then the definition of $\omega(T_v,i+1)$ implies that $\omega(T_v,i) = \omega(f) \geq \omega(T_v, i+1)$. One may thus assume that $c_{i+1} \in C(f,v)$. Then, interchanging colors $c_i$ and $c_{i+1}$ in $P(v;c_i,c_{i+1})$, we obtain another edge-coloring $f'$ of $T$. By Lemma [1] $\omega(f) \geq \omega(f')$ and $c_{i+1} \in \text{Miss}(f',v)$. Therefore by the definition we have

$$\omega(T_v,i) = \omega(f) \geq \omega(f') \geq \omega(T_v,i+1).$$

Q.E.D.

**Lemma 3.** For any vertex $v$ of tree $T$,

$$\omega(T_v,d(v)+1) = \omega(T_v,d(v)+2) = \cdots = \omega(T_v,n).$$

**Proof.** Let $i$ be any index such that $d(v) + 2 \leq i \leq n$. Then by Lemma [2] $\omega(T_v,d(v)+1) \geq \omega(T_v,i)$. It therefore suffices to prove that $\omega(T_v,d(v)+1) \leq \omega(T_v,i)$.
Let $f$ be any edge-coloring of $T_v$ such that $c_i \in \text{Miss}(f, v)$ and $\omega(f) = \omega(T_v, i)$. Since $|C(f, v)| = d(v)$, at least one of the $d(v)+1$ colors $c_1, c_2, \ldots, c_{d(v)+1}$ is missing at $v$. Let $c_j$ be any of them, then $1 \leq j \leq d(v)+1$ and $c_j \in \text{Miss}(f, v)$. By the definition of $\omega(T_v, j)$ we have

$$\omega(T_v, j) \leq \omega(f) = \omega(T_v, i).$$

(1)

Since $j \leq d(v)+1$, by Lemma 2

$$\omega(T_v, d(v)+1) \leq \omega(T_v, j).$$

(2)

Thus by Eqs. (1) and (2) we have $\omega(T_v, d(v)+1) \leq \omega(T_v, i)$.  

Q.E.D.

From Lemmas 2 and 3 we immediately have the following lemma.

**Lemma 4.** For any vertex $v$, $\omega(T_v) = \omega(T_v, d(v)+1)$. In particular, $\omega(T) = \omega(T_r, d(r)+1)$.

Thus our DP table for each vertex $v$ consists of only the $d(v)+1$ values $\omega(T_v, i)$, $1 \leq i \leq d(v)+1$, and hence the size is $d(v)+1$. All the values $\omega(T_v, i)$ for $i$, $d(v)+2 \leq i \leq n$, are equal to $\omega(T_v, d(v)+1)$.

For each leaf $v$ of $T$, the DP table for $v$ consists of a single value $\omega(T_v, 1) = 0$.

Note that $d(v) = 0$ and $T_v$ has no edge.

Thus it suffices to show how to construct the DP table for an internal vertex $v$ from the DP tables for all the children $v_1, v_2, \ldots, v_{d(v)}$ of $v$. (See Figure 2) We first notice that, for each $i$, $1 \leq i \leq d(v)+1$, $T_v$ has an edge-coloring $f$ such that $\omega(f) = \omega(T_v, i)$, $c_i \in \text{Miss}(f, v)$, and $C(f, v)$ consists of the first $d(v)$ colors in $C$ other than $c_i$, as in the following lemma whose proof is omitted in this extended abstract.

**Lemma 5.** For any vertex $v$ of $T$ and any index $i$, $1 \leq i \leq d(v)+1$, $T_v$ has an edge-coloring $f$ such that $\omega(f) = \omega(T_v, i)$, $c_i \in \text{Miss}(f, v)$, and $C(f, v) = D_i$, where $D_i = \{c_k \mid 1 \leq k \leq d(v)+1, k \neq i\}$.

We then compute the value $\omega(T_v, i)$ for an index $i$, $1 \leq i \leq d(v)+1$, from the $d(v) \times (d(v)+1)$ values $\omega(T_{v_j}, k)$ for all indices $j$, $1 \leq j \leq d(v)$, and all indices $k$, $1 \leq k \leq d(v)+1$. Let $J$ be the set of indices $j$ of all children $v_j$ of $v$, that is, $J = \{j \mid 1 \leq j \leq d(v)\}$. Let $e_j = (v, v_j)$ for each $j \in J$, as illustrated in Figure 2. Let $1 \leq i \leq d(v)+1$, and let $K_i$ be the set of indices $k$ of all colors $c_k$ in $D_i$, that is, $K_i = \{k \mid 1 \leq k \leq d(v)+1, k \neq i\}$.

Suppose that $f$ is an edge-coloring of $T_v$ such that $\omega(f) = \omega(T_v, i)$, $c_i \in \text{Miss}(f, v)$, and $C(f, v) = D_i$. Lemma 5 implies that $T_v$ has such an edge-coloring $f$. For each $j \in J$, let $f_j$ be the restriction of $f$ to the edges in $T_{v_j}$, that is, $f_j$ is an edge-coloring of $T_{v_j}$ such that $f_j(e) = f(e)$ for every edge $e$ in $T_{v_j}$. Then clearly

$$\omega(f) = \omega(T_v, i) = \sum_{j \in J} \omega(f_j) + \omega(D_i),$$

(3)
where
\[ \omega(D_i) = \sum_{c \in D_i} \omega(c). \]

Note that the value \( \omega(D_i) \) depends on \( i \), but does not depend on the edge-coloring \( f \). For each \( j \in J \), we denote by \( j' \) the index such that \( c_{j'} = f(e_j) \in D_i \). Since \( f \) is an edge-coloring of \( T_v \), all colors \( f(e_j) \) are different from each other. Therefore the mapping \( \varphi : J \to K_i \) such that \( \varphi(j) = j' \in K_i \) for each \( j \in J \) is a bijection. Since \( c_{j'} \in \text{Miss}(f_j, v_j) \), the definition of \( \omega(T_{v_j}, j') \) implies that \( \omega(f_j) \geq \omega(T_{v_j}, j') \). However, we have
\[ \omega(f_j) = \omega(T_{v_j}, j') \] (4)
as follows: if \( \omega(f_j) > \omega(T_{v_j}, j') \), then \( T_{v_j} \) has an edge-coloring \( f'_j \) such that \( c_{j'} \in \text{Miss}(f'_j, v_j) \) and \( \omega(f'_j) = \omega(T_{v_j}, j') < \omega(f_j) \), and hence the edge-coloring \( f' \) of \( T_v \) obtained from \( f \) by replacing \( f_j \) with \( f'_j \) satisfies \( c_i \in \text{Miss}(f', v) \) and \( C(f', v) = D_i \) but \( \omega(f') < \omega(f) = \omega(T_v, i) \), a contradiction. By Eqs. (3) and (4) we have
\[ \omega(f) = \omega(T_v, i) = \sum_{j \in J} \omega(T_{v_j}, j') + \omega(D_i). \] (5)

Suppose conversely that \( \varphi : J \to K_i \) is any bijection. For each \( j \in J \), we denote \( \varphi(j) \) by \( j' \), that is, \( j' = \varphi(j) \), and let \( g_j \) be any edge-coloring of \( T_{v_j} \) such that \( \omega(g_j) = \omega(T_{v_j}, j') \) and \( c_{j'} \in \text{Miss}(g_j, v_j) \). Let \( g \) be an edge-coloring of \( T_v \) extended from all \( g_j, j \in J \), as follows:
\[ g(e) = \begin{cases} c_{j'} & \text{if } e = (v, v_j) \text{ and } j \in J; \\ g_j(e) & \text{if } e \text{ is an edge of } T_{v_j} \text{ and } j \in J. \end{cases} \]

Then \( c_i \in \text{Miss}(g, v) \), \( C(g, v) = D_i \), and
\[ \omega(g) = \sum_{j \in J} \omega(g_j) + \omega(D_i) = \sum_{j \in J} \omega(T_{v_j}, j') + \omega(D_i). \] (6)

Let
\[ b(i) = \min_{\varphi} \sum_{j \in J} \omega(T_{v_j}, j'), \] (7)
where the minimum is taken over all bijections \( \varphi : J \to K_i \). Then one can know from Eqs. (5) and (6) that \( \omega(g) = \omega(T_v, i) \) if and only if \( \varphi \) attains the minimum value \( b(i) \).

We thus have the following lemma.

**Lemma 6.** For any internal vertex \( v \) of \( T \) and any index \( i, 1 \leq i \leq d(v) + 1 \),
\[ \omega(T_v, i) = b(i) + \omega(D_i). \]
One can easily compute the values $\omega(D_i)$, $1 \leq i \leq d(v) + 1$, in time $O(d(v))$. We shall therefore show how to compute $b(i)$.

The problem of computing $b(i)$ can be reduced to the minimum weight perfect matching problem for a complete bipartite graph, as follows. Let $B(i) = (J, K_i; J \times K_i)$ be a complete bipartite graph with vertex sets $J$ and $K_i$. (See Figure 3) Note that $|J| = |K_i| = d(v)$. Let $h : J \times K_i \rightarrow \mathbb{R}$ be a weight function such that $h(e) = \omega(T_{v,j}, k)$ for each edge $e = (j, k)$, $j \in J$ and $k \in K_i$. The minimum weight perfect matching problem is to find a perfect matching $M$ such that the sum of weights of the edges in $M$ is as small as possible [12]. A perfect matching in $B(i)$ is indicated by thick lines in Figure 3. Clearly, $b(i)$ is equal to the weight of the minimum perfect matching in $B(i)$.

The minimum weight perfect matching problem can be solved in time $O(d^3(v))$ for a single graph $B(i)$ [12], and hence the value $b(i)$ for an index $i$, $1 \leq i \leq d(v) + 1$, can be computed in time $O(d^3(v))$. Therefore the $d(v) + 1$ values $b(i)$ for all indices $i$, $1 \leq i \leq d(v) + 1$, can be computed total in time $O(d^4(v))$.

The third idea is to compute all the $d(v) + 1$ values $b(i)$, $1 \leq i \leq d(v) + 1$, together. Kao et al. recently showed that if a given complete bipartite graph $B$ has $d$ vertices then one can solve in time $O(d^3)$ the minimum weight perfect matching problems for all $d$ complete bipartite graphs obtained from $B$ by deleting each of the $d$ vertices [7,8]. Let $K = \{1, 2, \ldots, d(v) + 1\}$, let $B = (J, K; J \times K)$ be a complete bipartite graph, and let $h(e) = \omega(T_{v,j}, k)$ for each edge $e = (j, k)$, $j \in J$ and $k \in K$. Applying their algorithm to $B$, we can compute all the values $b(i)$, $1 \leq i \leq d(v) + 1$, in time $O(d^3(v))$. We thus have the following lemma.

**Lemma 7.** For any internal vertex $v$ of $T$, one can construct the DP table for $v$ from the DP tables for all the children of $v$ in time $O(d^3(v))$.

We can thus compute the minimum cost $\omega(T) = \omega(T_r, d(r) + 1)$ in the DP table for root $r$ in time.
\[
O \left( \sum_{v \in V} d^3(v) \right) = O \left( \left( \sum_{v \in V} d(v) \right) \Delta^2 \right) \\
= O(n\Delta^2).
\]

Clearly, the total size of all DP tables is \(\sum_{v \in V} (d(v) + 1) = O(n)\). We do not store the bijection \(\varphi\) found at each internal vertex.

### 3.2 Finding an Optimal Edge-Coloring

We next show how to find an optimal edge-coloring \(f\) of \(T\). Although we compute the DP tables from leaves to the root, we decide the colors of edges from the root to leaves, as follows.

When we computed

\[
\omega(T) = \omega(T_r, d(r) + 1) = b(d(r) + 1) + \omega(D_{d(r)+1}),
\]

we found a bijection \(\varphi : J \to K_{d(r)+1}\) attaining the minimum value \(b(d(r) + 1)\) in Eq. (7), where \(J = K_{d(r)+1} = \{j \mid 1 \leq j \leq d(r)\}\). We find the bijection \(\varphi\) again. For each \(j, 1 \leq j \leq d(r)\), we color the edge \(e_j\) joining \(r\) and its \(j\)th child \(r_j\) with \(c_{\varphi(j)} \in D_{d(r)+1} = \{c_1, c_2, \ldots, c_{d(r)}\}\). We hence have \(C(f, r) = \{c_1, c_2, \ldots, c_{\text{deg}(r)}\}\) since \(d(r) = \text{deg}(r)\).

We may thus assume that, for an internal vertex \(v\), the edge joining \(v\) and its parent \(u\) has been colored with a color \(c_i \in C\) for some index \(i, 1 \leq i \leq \text{deg}(u)\). Let \(i' = \text{deg}(u) + 1\) if \(i \leq \text{deg}(u)\), and let \(i' = \text{deg}(u) + 1\) if \(i > \text{deg}(u)\). We find again the bijection \(\varphi : J \to K_{i'}\) attaining the minimum value \(b(i')\) in Eq. (7). For each \(j, 1 \leq j \leq d(v)\), we color the edge \(e_j = (v, v_j)\) with \(c_{\varphi(j)} \in D_{i'}\). We hence have \(C(f, v) = D_{i'} \cup \{c_i\} \subseteq \{c_1, c_2, \ldots, c_{\text{deg}(v)}\} \cup \{c_i\}\) since \(d(v) + 1 = \text{deg}(v)\).

Thus we can correctly find an optimal edge-coloring of \(T\). Clearly, the algorithm takes time \(O(n\Delta^2)\). This completes a proof of Theorem 1.

Kao et al. also show that if a given complete bipartite graph \(B\) has \(d\) vertices and all edge weights are integers in the range \([-N, N]\) then one can solve in time \(O(d^{2.5}\log(dN))\) the minimum weight perfect matching problems for all \(d\) complete bipartite graphs obtained from \(B\) by deleting each of the \(d\) vertices \([7, 8]\).

Using their algorithm, one can find an optimal edge-coloring of tree \(T\) in time \(O(n\Delta^{1.5}\log(\Delta N))\) if all the color costs are integers in the range \([-N, N]\).

We finally give a remark. As we mentioned in Introduction, an optimal edge-coloring of a graph \(G\) does not always use the minimum number \(\chi'(G)\) of colors. However, from the algorithm we immediately have the following theorem for trees \(T\). Note that \(\chi'(T) = \Delta(T)\) and that our algorithm does not use the method of interchanging colors in an alternating path although it is used in the proofs of Lemmas 2 and 5.

**Theorem 2.** Any tree \(T\) has an optimal edge-coloring using the first \(\Delta(T)\) cheapest colors \(c_1, c_2, \ldots, c_\Delta\) in \(C\).
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References

Counting $H$-Colorings of Partial $k$-Trees*

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Abstract. The problem of counting all $H$-colorings of a graph $G$ of $n$ vertices is considered. While the problem is, in general, #P-complete, we give linear time algorithms that solve the main variants of this problem when the input graph $G$ is a $k$-tree or, in the case where $G$ is directed, when the underlying graph of $G$ is a $k$-tree. Our algorithms remain polynomial even in the case where $k = O(\log n)$ or in the case where the size of $H$ is $O(n)$. Our results are easy to implement and imply the existence of polynomial time algorithms for a series of problems on partial $k$-trees such as core checking and chromatic polynomial computation.

1 Introduction

In this paper we consider a series of counting problems associated with various versions of homomorphisms from a graph $G$ to a fixed graph $H$. Given a graph $H = (V(H), E(V))$, with loops but without multiple edges, for any graph $G = (V(G), E(G))$ an homomorphism of $G$ to $H$ is a map $\theta : V(G) \rightarrow V(H)$ with the property that $\{v, w\} \in E(G) \Rightarrow \{\theta(v), \theta(w)\} \in E(H)$. We will call such a homomorphism a $H$-coloring of $G$. For $H$ fixed, the $H$-coloring problem asks if there is an $H$-coloring of $G$. If $H$ is bipartite or it has a loop, then the $H$-coloring problem can be trivially solved in polynomial time, but in the case that $H$ is not-bipartite and loopless the problem is known to be NP-complete [HN90] (see also [GHN00] for the complexity of the same problem for bounded degree graphs).

An earlier version of the $H$-coloring problem is the exact $H$-coloring problem: given a graph $G$, decide if there is an $H$-coloring $\theta$ of $G$ such that $\theta(V(G)) = V(H)$ and $\theta(E(G)) = E(H)$. This problem appeared in [GJ79] as problem GT-52 and is known to be NP-complete, even for the case where $H$ is a triangle [Lev73]. We also consider the intermediate problem of asking whether there is an $H$-coloring $\theta$ of $G$ such that $\theta(V(G)) = V(H)$ and we call it vertex exact $H$-coloring.

Given a graph $G$, let $\mathcal{H}(G, H)$ denote the set of all $H$-colorings of $G$, and let $\mathcal{E}(G, H)$ denote the set of all exact $H$-colorings of $G$. Also, we denote as

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The set of all vertex $H$-colorings of $G$. For a fixed graph $H$, denote by $\#H$-coloring the problem of, given $G$ as input, computing $|\mathcal{H}(G, H)|$. Also, denote by $\#E H$-coloring the problem of, given as input $G$, computing $|\mathcal{E}(G, H)|$ and by $\#VE H$-coloring the problem of, given as input $G$, computing $|\mathcal{V}(G, H)|$. Dyer and Greehill have proved that unless every connected component of $H$ is an isolated vertex without a loop, a complete graph with all loops present or a complete unlooped bipartite graph, the $\#H$-coloring problem is $\#P$-complete, even when $G$ has bounded degree. It is easy to see that the same result hold for the $\#E H$-coloring ($\#V E H$-coloring) as the knowledge of $\mathcal{E}(G, H')$ ($\mathcal{V}(G, H')$) for any (induced) subgraph $H'$ of $H$ implies the knowledge of $\mathcal{H}(G, H)$. For further negative results on the approximability of the $\#H$-coloring see [CDF] and [DFJ99]. For positive and negative complexity results concerning special cases of the $\#H$-coloring problem, depending on restrictions on both $H$ and $G$, see [Edw86, DNS00].

The notion of treewidth appears to play a central role in many areas of graph theory. Roughly, a graph has small treewidth if it can be constructed by assembling small graphs together in a tree structure, namely a tree decomposition of small width (see section 2 for the formal definitions). It was first defined by Robertson and Seymour and served as one of the cornerstones of their lengthy proof of the Wagner conjecture, known now as the Graph Minors Theorem (for a survey see [RS85]). Treewidth appears to have interesting applications in algorithmic graph theory. In particular, a wide range of, otherwise intractable, combinatorial problems are polynomially, even linearly, solvable when restricted to graphs with bounded treewidth (alternatively, graphs with bounded treewidth are called partial $k$-trees). Several general results have been developed for proving, constructively, the existence of such algorithms. Among them, we mention the work of Courcelle in [Cou91b] where the existence of polynomial algorithms is guarantied by the expressability of the corresponding problem by Monadic Second Order formulas (see also [Cou90a]). Similar results have been developed in [CM98] for solving counting problems for partial $k$-trees (see also [CMR]). As a consequence of these results, it is possible to contruct a polynomial time algorithm solving the $\#H$-coloring problem for partial $k$-trees when $k$ and the size of $H$ are fixed constants. Unfortunately, the general results of Courcelle, do not provide implementable algorithms because of the immense hidden constants in their complexity. Alternatively, the attention of many researchers was fixed on developing fast and easy to implement “tailor made” algorithms for problems of specific interest. The standard methodology to get such solutions consists of a two step procedure: First to find a tree-decomposition of the input graph, with treewidth bounded by a constant, although possibly not optimal, and second to use some kind of dynamic programming taking advantage of the bounded treewidth decomposition of the graph, to get a solution (for a survey, see [Bod97]). In particular, such a dynamic programing technique, solving – among others – the $H$-coloring problem for partial $k$-trees, has been developed in [TP97] (also, see [JS97], for counting list colorings and other generalizations of colorings). For the first canonical step, Bodlaender in [Bod96] presented, for
any $k$, a linear time algorithm that, given a graph $G$, checks whether $G$ is a partial $k$-tree and, if so, outputs an minimum width tree decomposition. As this algorithm appears extremely hard to be implemented (see also [BK96]), the first canonical step can be carried out by one of the previously developoed algorithms that are able to output a tree-decomposition of $G$ with treewidth bounded by a constant (linear in $k$) [Ree92]. In particular, the deterministic algorithm by Reed [Ree92] runs in $O(n \log n)$ time, which, for constant $k$, gives a 2 approximation to the optimal treewidth-decomposition. Moreover, its implementation is easier than the algorithm in [Bod96]. Consequently, we will assume that any partial $k$-tree is always accompanied with its linear size tree decomposition of bounded width.

In this paper we present a polynomial time algorithm for the counting problems $\#H$-coloring, $\#EH$-coloring, and $\#VEH$-coloring, in the case that the input graph $G$ has constant treewidth. Although our methodology follows the two cannonical steps, we shall remark that we present an easy to be implemented algorithm for the second step, assuming the aforementioned results on the existence of fast and implementable algorithms for obtaining a bounded treewidth-decomposition for a partial $k$-tree $G$. Finally, we stress out that our results are the first “tailor made” positive results on counting $H$-colorings for the case where $H$ is generic and $G$ is restricted.

In Section 2 we review the basic definitions and some preliminary results on treewith. In Section 3 we present the algorithm solving the $\#H$-coloring problem. In Section 4 we present two variants of the algorithm of Section 3, for solving the $\#VEH$-coloring and the $\#EH$-coloring problem respectively. In the final section, we present some consequences of our previous results for a series of problems on graphs with bounded treewidth.

## 2 Definitions and Basic Results

Robertson and Seymour introduced the notion of treewidth in the framework of the theory of graph minors.

A tree decomposition of a graph $G$ is a pair $(X, U)$ where $U$ is a tree whose vertices we will call nodes and $X = \{X_i \mid i \in V(U)\}$ is a collection of subsets of $V(G)$ such that

1. $\bigcup_{i \in V(U)} X_i = V(G)$,
2. for each edge $\{v, w\} \in E(G)$, there is an $i \in V(U)$ such that $v, w \in X_i$, and
3. for each $v \in V(G)$ the set of nodes $\{i \mid v \in X_i\}$ forms a subtree of $U$.

The width of a tree decomposition $(\{X_i \mid i \in V(U)\}, U)$ equals $\max_{i \in V(U)} \{|X_i| - 1\}$. The treewidth of a graph $G$ is the minimum width over all tree decompositions of $G$. Computing treewidth is an NP-complete problem [ACP93].

We now define some extensions of the definition of a tree decomposition were introduced by Bodlaender and Kloks [Bod96] (see also [Arn85]). A rooted tree decomposition is a triple $D = (X, U, r)$ in which $U$ is a tree rooted at $r$ and $(X, U)$ is a tree decomposition. Let $D = (X, U, r)$ be a rooted tree decomposition
of a graph $G$ where $X = \{X_i \mid i \in V(U)\}$. $D$ is called a nice tree decomposition if the following are satisfied:

1. Every node of $U$ has at most two children,
2. if a node $i$ has two children $j,h$ then $X_i = X_j = X_h$,
3. if a node $i$ has one child, then either $|X_i| = |X_j| + 1$ and $X_j \subseteq X_i$ or $|X_i| = |X_j| - 1$ and $X_i \subseteq X_j$.

The following lemma can be found in [BK96].

**Lemma 1.** For any constant $k \geq 1$, given a tree decomposition of a graph $G$ of width $\leq k$ and $O(n)$ nodes, where $n$ is the number of vertices in $G$, there exists an $O(n)$ algorithm that constructs a nice tree decomposition of $G$ of width less or equal than $k$ and with at most $O(n)$ nodes.

We observe that a nice tree decomposition $D = (\{X_p \mid p \in V(U)\}, U, r)$ contains nodes of the following four possible types. A node $p \in V(U)$ is called: Start, if $p$ is a leaf, Join, if $p$ has two children $q_i,i = 1,2$, Forget, if $p$ has only one child $q$ and $|X_p| = |X_q| - 1$, Introduce, if $p$ has only one child $q$ and $|X_p| = |X_q| + 1$.

Notice that for every start node we may assume that $|X_p| = 1$: the effect of start nodes with $|X_p| \geq 2$ can be obtained by using a start node with a set containing one vertex, and then $|X_p| - 1$ introduce nodes, which add all the other vertices. For the purposes of this paper we will assume that the root $r$ is a forget node with $X_r = \emptyset$. Taking in mind Lemma 1 we will assume that any partial $k$-tree is given along with a nice tree decomposition of it.

Let $D = (X, U, r)$ be a nice tree decomposition of a graph $G$. For each node $p$ of $D$, let $U_p$ be the subtree of $U$, rooted at node $p$. We set $V_p = \cup_{v \in V(U_p)} X_v$ and, for any $p \in V(U)$, we define $G_p = G[V_p]$. Notice that $G_r = G$.

## 3 An Algorithm for Counting the Number of $H$-Colorings

Let $H$ be a fixed graph. Given an input graph $G$ together with a nice tree decomposition $D = (X, U, r)$ of it, we wish to count $|\mathcal{H}(G,H)|$, i.e. to solve the #H-coloring problem for $G$.

The algorithm will do a traversal of $D$ computing and tabulating for each $p \in V(U)$ the appropriate information. Before describing it, we need some additional definitions.

Given a nice tree decomposition $D = (X, U, r)$ of a graph $G$, for each $p \in V(U)$, let $F_p = \{\varphi : X_p \to V(H)\}$. Notice that if $G$ has treewidth $k$, then for any $p \in V(U)$, $|F_p| \leq h^{k+1}$. Moreover, for the root we have $F_r = \{\emptyset\}$.

The table associated to a node $p \in V(U)$ will have an entry for each $\varphi \in F_p$, holding the value $I_p(\varphi) = |\{\theta \in \mathcal{H}(G_p,H) \mid \theta|_{X_p} = \varphi\}|$ (in general, if $\theta$ is a function, we denote $\theta|_{S} = \{(v,a) \in \theta \mid v \in S\}$). As we always have $\theta|_{\emptyset} = \emptyset$, we get $|\mathcal{H}(G,H)| = I_r(\emptyset)$.

The function $\text{count-}H$ takes as input $G$, $D$ and a topological ordering $S = (p_1, \ldots, p_s)$ of $V(U)$ and outputs the value $I_r(\emptyset)$. 
function count-H(G, H, D, S)

\[ D = (X, U, r) \] is a nice tree decomposition of \( G \) with width \( k \)
\[ n = |V(G)|, \ h = |V(H)|, \ s = |V(U)| \]
\[ S = (p_1, \ldots, p_s) \] is a topological ordering of \( V(U) \)
begin
    for \( p := p_1 \) to \( p_s \) do
        if \( p \) is a start node, with \( X_p = \{ v \} \) then
            for all \( a \in V(H) \) do \( I_p((v, a)) := 1 \)
        end
        if \( p \) is a introduce node then
            let \( q \) be its unique child
            \( v := X_p - X_q \)
            \( S_q := \{ u \in X_q \mid \{ u, v \} \in E(G_p) \} \)
            for all \( \varphi \in F_q \) and \( a \in V(H) \) do
                if \( \forall u \in S_q \{ \varphi(u), a \} \in E(H) \) then
                    \( I_p(\varphi \cup \{ (v, a) \}) := I_q(\varphi) \)
                else \( I_p(\varphi \cup \{ (v, a) \}) := 0 \)
            end
        end
        erase the information on node \( q \)
    if \( p \) is a forget node then
        let \( q \) be its unique child
        \( v := X_q - X_p \)
        for all \( \varphi \in F_p \) do
            \( I_p(\varphi) := \sum_{a \in V(H)} I_q(\varphi \cup \{ (v, a) \}) \)
        end
        erase the information on node \( q \)
    end
    if \( p \) is a join node with children \( q_1 \) and \( q_2 \) then
        for all \( \varphi \in F_p \) do
            \( I_p(\varphi) := I_{q_1}(\varphi) \cdot I_{q_2}(\varphi) \)
        end
        erase the information on nodes \( q_1 \) and \( q_2 \)
    end
    end
return \( I_r(\varnothing) \)
end

Theorem 1. Given a graph \( G \) with \( n = |V(G)| \), a nice tree decomposition \( D = (X, U, r) \) of \( G \) with width \( k \), and a topological ordering of the nodes in \( D \). Then there exists an algorithm that computes the number of \( H \)-colorings of \( G \) in \( O(nh^{k+1} \min\{k, h\}) \) steps using \( O(h^{k+1} \log n) \) additional space, where \( h = |V(H)| \).

Notice that for constant treewidth \( k \), we can allow \( h = O(n) \) and still have a polynomial time algorithm. Taking in mind the algorithm of Bodlaender in [Bod96] we have the following.

Corollary 1. Given a graph \( G \) if it has treewidth bounded by a constant \( k \), then we can compute \( |\mathcal{H}(G, H)| \) in \( O(nh^{k+1} \min\{k, h\}) + O(f(k)n) \) steps.
For constant $k$ and $h$, our algorithm works in linear time and therefore we have the following.

**Corollary 2.** The \#H-coloring problem, with parameters the treewidth of the input and the size of $H$, is a fixed parameter tractable problem (i.e. belongs to the class FPT).

By suitably altering function $\text{count-H}$ we can have the analogous of Theorem 1 and Corollary 1 for the \#VEH-coloring and the \#EH-coloring problems.

**Theorem 2.** Given a graph $G$ with $n = |V(G)|$, a nice tree decomposition $D = (X, U, r)$ of $G$ with width $k$, and a topological ordering of the nodes in $D$. Then there exists an algorithm that computes the number of vertex exact $H$-colorings of $G$ in $O(nh^k 2^h \max\{k, 2^{2h}\})$ time using $O(h^k 2^h \log n)$ additional space, where $h = |V(H)|$.

**Corollary 3.** Given a graph $G$ with $n$ vertices and treewidth bounded by a constant $k$, then we can compute $|V(G, H)|$ in $O(nh^k 2^h \max\{k, 2^{2h}\}) + O(f(k)n)$ time, where $h = |V(H)|$.

**Theorem 3.** Given a graph $G$ with $n = |V(G)|$, a nice tree decomposition $D = (X, U, r)$ of $G$ with width $k$, and a topological ordering of the nodes in $D$. Then there exists an algorithm that computes the number of $H$-colorings of $G$ in $O(nh^k 2^{h+e} \max\{k, 2^{2(h+e)}\})$ time using $O(h^k 2^{h+e} \log n)$ additional space, where $h = |V(H)|$ and $e = |E(H)|$.

**Corollary 4.** Given a graph $G$ with $n$ vertices and treewidth bounded by a constant $k$, then we can compute $|E(G, H)|$ in $O(nh^k 2^{h+e} \max\{k, 2^{2(h+e)}\}) + O(f(k)n)$ time, where $h = |V(H)|$ and $e = |E(H)|$.

Notice that for each of the algorithms given so far, if we retain the information of all the tables, it is possible to use it in order to enumerate all the homomorphisms. The storing of all the tables implies a burden of $O(n/\log n)$ to the space reported in Theorems 1, 2, and 3. In particular, setting up a suitable bookkeeping of the enumerated homomorphisms, a top-down traversal of the table information can pop-up each of them in $O(n)$ steps.

We point out that in all algorithms of this section and for constant $h$, we can allow $k = O(\log n)$ and still have polynomial time. Finally, all the results of this section can be adapted straightforwardly to solve the directed versions of the problems ($H$ and $G$ are directed graphs).

### 4 Extensions to Other Results

In this section, we present some consequences of the polynomial time algorithms given in the previous sections. For each problem, either we give a new polynomial-time result or we improve previously known results. Recall that if we are given a graph $G$, directed or undirected, all of our algorithms should be used in conjunction with one of the algorithms mentioned in the introduction, that in polynomial time, compute the treewidth decomposition of bounded width.
4.1 Coloring Problems

Notice that a graph \( G \) is \( H \)-colorable if \( |\mathcal{H}(G,H)| \geq 1 \). This implies the following corollary of Theorems 1, 2, and 3 for the corresponding decision version of the problems examined.

**Corollary 5.** For any graph \( H \) with \( h \) vertices and \( e \) edges, we can solve, for an \( n \)-vertex partial \( k \)-tree as input graph,

- the \( H \)-coloring problem in \( O(nh^{k+1}\min\{k,h\}) \) steps,
- the vertex exact \( H \)-coloring problem in \( O(nh^{k+1}2^h\max\{k,2^h\}) \) steps, and
- the exact \( H \)-coloring problem in \( O(nh^{k+1}2^h+e\max\{k,2^h+e\}) \) steps.

Corollary 5 improves the time of the best known algorithm for the \( H \)-coloring of partial \( k \)-trees. This algorithm runs in \( O(nh^{2(k+1)}) \) steps and is due to Telle and Proskurowski in [TP97].

We denote as \( K_c \) the complete graph with \( c \) vertices. As the problem of counting the number of \( c \)-colorings of a graph \( G \) is equivalent to the \#\( K_c \)-coloring problem we can conclude to the following.

**Corollary 6.** An algorithm can be constructed to compute the number of \( c \)-colorings of a partial \( k \)-tree of \( n \) vertices in \( O(nc^{k+1}\min\{k,c\}) \) steps.

We denote as \( \beta_G(c) \) the function mapping \( c \) to the number of colorings of \( G \) that use exactly \( c \) colors, i.e. \( \beta_G(c) = |\mathcal{V}(G,K_c)| \).

**Lemma 2.** For any graph \( G \), \( \beta_G(c) = |\mathcal{H}(G,K_c)| - \sum_{1 \leq r \leq c-1} \binom{c}{r} \beta_G(r) \).

The formula of Lemma 2 and Corollary 6 give a way to compute \( \beta_G(c) \).

**Lemma 3.** An algorithm can be constructed that, given partial \( k \)-tree \( G \) with \( n \) vertices and for any \( c,1 \leq c \leq n \), computes \( \beta_G(c) \) in \( O(nc^{k+2}\min\{k,c\}) \) steps.

**Corollary 7.** An algorithm can be constructed that computes the chromatic polynomial of a partial \( k \)-tree with \( n \) vertices in \( O(kn^{k+3}) \) steps.

In our knowledge, the best algorithm for computing the chromatic polynomial for partial \( k \)-trees can be derived from the algorithm of Andrzejak [And98]. This algorithm can compute in \( O(n^{2+7\log_2 c}) \) steps the Tutte polynomial of a partial \( k \)-tree of \( n \) vertices, where \( c \) is twice the number of partitions of a set with \( 3(k+1) \) elements. Several researchers have considered the problem of counting the number of proper \( c \)-colorings in a graph \( G \). The problem is \#P-hard for \( c \geq 3 \) and maximum degree \( \Delta \geq 3 \) [Jer95]. In the same paper, it is proved that there exists a Fully Polynomial time Randomized Aproximation Scheme (FPRAS) for the number of colorings in the case when \( c \geq 2\Delta + 1 \). Recently, Bubley et all. in [BDGJ99] proved that the problem is \#P-hard for fixed \( \Delta \), but there is a FPRAS for \( c = 5 \) and \( \Delta = 3 \). Edwards [Edw89] proved that if \( c \geq 3 \) and the minimum degree where \( \delta \geq \alpha n \), \( n = |\mathcal{V}(G)| \), the counting problem is \#P-complete if \( \alpha < \frac{c-2}{c-1} \), but it is in \( P \) for \( \alpha > \frac{c-2}{c-1} \).
4.2 Problems on Cores

Given a graph $G$, a core of $G$ is a subgraph $C$ of $G$ such that $G$ is homomorphic to $C$, but $G$ fails to be homomorphic to any proper subgraph of $C$. This notion of core is due to Hell and Nešetřil [HN92]. A graph $G$ is a core if $G$ is a core for itself or, equivalently, if $|E(G,G)| = |H(G,G)|$. For example, any graph $C_{2k+1}$ is a core ($C_{2k+1}$ is the connected 2-regular graph with $2k+1$ vertices). It is known, that in general, the problem of deciding whether $G$ is a core is NP-complete [HN92] and in the same paper it is presented a polynomial time algorithm, to decide if $G$ is a core, for the particular case when $G$ has independence number $\alpha(G) \leq 2$. Finally, we mention that “almost all graphs” are cores [Neš99].

We first prove the following.

**Corollary 8.** An algorithm can be constructed that decides, in $O(k^2n^{k+3})$ steps, whether a partial $k$-tree with $n$ vertices is a core.

**Proof.** Notice that it is enough to check whether there exists a subgraph $H$ of $G$ where $|E(H)| = |E(G)| - 1$ and such that $|H(G,H)| \geq 1$. According to Theorem 1, this requires $O(|E(G)|kn^{k+2}) = O(k^2n^{k+3})$ steps (take in mind that if $G$ is a partial $k$-tree then $E(G) = O(k|V(G)|)$).

We mention that the property of being a core is an EMS-property (i.e. involves counting or summing evaluations over sets definable on monadic second order logic) and therefore, the results in [Cou90b, ALS91] imply the existence of a polynomial time algorithm for deciding it. So far, no explicit algorithm has been reported for this problem.

We consider the isomorphism problem on cores. Notice that graphs $G$ and $H$ are isomorphic iff $|E(G,H)| > 1$. This check can be done in polynomial time when both $G$ and $H$ are cores and partial $k$-trees.

**Corollary 9.** An algorithm can be constructed that checks, in $O(kn^{k+2})$ steps, if two cores $G$ and $H$ of treewidth $\leq k$ are isomorphic ($n = \max\{|V(G)|, |V(H)|\}$).

In general, checking whether two partial $k$-trees are isomorphic can be done in $O(n^{k+4.5})$ steps [Bod90]. Corollaries 8 and 9 improve this time for the case where some of the input graphs is a core.

Using the idea of Corollary 9, we can count automorphisms of cores with bounded treewidth.

**Corollary 10.** An algorithm can be constructed that outputs, in $O(kn^{k+2})$ steps, the number of automorphisms of a core $G$ of $n$ vertices and treewidth $\leq k$.

A graph $G$ is called rigid when the identity is the only homomorphism in $\mathcal{H}(G,G)$ [Neš99]. We conclude to the following.

**Corollary 11.** An algorithm can be constructed that checks, in $O(kn^{k+2})$, steps whether a $n$-vertex partial $k$-tree is rigid or not.
4.3 Problems on Counting Independent Sets

In the case where $H$ is the unique graph with two vertices one edge and one loop, $|H(G, H)|$ is the number of independent sets of $G$. We have the following.

**Corollary 12.** An algorithm can be constructed that outputs the number of independent sets of a $n$-vertex partial $k$-tree in $O(n2^{k+1})$ steps.

Notice that the linear algorithm of Corollary 12 remains polynomial even if the treewidth of $G$ is bounded by $O(\log n)$. We could remark that, by specializing the structure of $H$, we can generate an arbitrary number of counting problems that are, in general, #P-complete. As an example we mention the problem of counting the $q$-particle Widom-Rowlinson configurations in bounded treewidth graphs [DG00].

Suppose that $H$ is a star with a loop attached on its unique internal vertex (a star is a tree of diameter 2). Then, asking whether $|V(G, H)| \geq 1$ is equivalent to ask whether $G$ has an independent set of size $\geq k$. Finally, Theorem 2 yields the following.

**Corollary 13.** An algorithm can be constructed that for any partial $k$-tree of $n$ vertices and any $r$, outputs the number of the independent sets of $G$ with at least $r$ vertices in $O(n(r + 1)^{k+1}2^{r+1} \max\{k, 2^{2(r+1)}\})$ steps.

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References


Counting $H$-Colorings of Partial $k$-Trees


A Linear Time Algorithm for Enumerating All the Minimum and Minimal Separators of a Chordal Graph

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Abstract. We give $O(m + n)$ algorithms for enumerating all minimum separators as well as minimal separators in a connected chordal graph. We give a tight upper bound, $n - \kappa(G) - 1$, for the number of minimal separators in a chordal graph. An important contribution of the paper is our characterisation of minimal (minimum) separators of chordal graphs and the above results are obtained as consequences of the characterisations.

1 Introduction

Let $C$ be a cycle in a graph $G$. A chord of $C$ is an edge of $G$ joining two vertices of $C$ which are not consecutive. A graph $G$ is called a chordal (or triangulated) graph iff every cycle in $G$ of length 4 or more has a chord. Chordal graphs arise in many applications (see [6,15,20]). Chordal graphs constitute one of the most important subclasses of perfect graphs [6].

In a connected graph $G$, a separator $S$ is a subset of vertices whose removal separates $G$ into at least two connected components. $S$ is called a $(a - b)$ separator iff it disconnects vertices $a$ and $b$. A $(a - b)$ separator is said to be a minimal separator iff it does not contain any other $(a - b)$ separator. A minimum separator is a separator with minimum size. Clearly every minimum separator is a minimal separator. The (vertex) connectivity, $\kappa(G)$ is defined to be the size of a minimum separator. The connectivity of a complete graph is by definition $|V| - 1$. It doesn’t have any separators. For basic notions about connectivity, we refer to [7,6].

The problem of listing all minimal separators and minimum separators is one of the fundamental enumeration problems in graph theory, which has great practical importance in reliability analysis for networks and operations research for scheduling problems [9,5,1].

The related problem of finding the connectivity of a graph is conjectured to have an $O(m + n)$ algorithm (see [4] for history). The interest in finding better or different algorithms is still very much alive [19,16,17,13,12]. The latest and the fastest in the sequence is due to Gabow [4], which runs in $O(n^2k + nk(min(k^2, kn^2)))$ time, where $k = \kappa(G)$. In this paper, we present an $O(m + n)$
algorithm, restricted to the class of chordal graphs, which not only finds the connectivity, but lists all minimum separators.

The problem of listing all minimum or minimal separators of an undirected graph is considered by various authors [5,9,10]. In [9], a $O(2^k n^3)$ algorithm is given for listing all minimum separators. A $O(n^6 R^2)$ algorithm is given in [10] to list all minimal separators, where $R$ is the total number of minimal separators in the graph. This is improved in [8] to $O(n^3 R^2 + n^4 R)$, where $R \leq (n(n-1)/2 - m) R$. Restricted to the class of chordal graphs, we present algorithms that lists all minimal (or minimum) separators in $O(m + n)$ time, which compares sharply with the general case running times available in [9,10,8]. Algorithms to enumerate the minimal separators for some other subclasses of perfect graphs (e.g., permutation graphs) are given in [12,11].

As an auxiliary result we show that the number of minimal separators in a chordal graph is upper bounded by $n - \kappa(G) - 1$. We also show the tightness of this upper bound by constructing a graph on $n$ nodes with number of minimum separators = number of minimal separators = $n - \kappa(G) - 1$. Some other important classes of graphs which are known to have polynomial number of minimal separators include permutation graphs, circular permutation graphs, trapezoidal graphs, circle graphs, circular arc graphs, distance hereditary graphs, chordal bipartite graphs, cocomparability graphs of bounded dimension and weakly triangulated graphs [11,12,13,14]. Note that in general graphs, the number of minimal separators can be exponential [10]. In [9] it is shown that in a general graph, the number of minimum separators is bounded by $O(2^{k_2} R^2)$, where $k = \kappa(G)$. For information on why number of minimal (or minimum) separators are important see [10,12].

An important contribution of this paper is the structural understanding of separators of chordal graphs in general and the minimum and minimal separators in particular (Lemma 5, Theorems 1, 2). As far as we know, these characterisations are new. Also we haven’t seen any algorithm for listing the minimal separators of a chordal graph elsewhere. But we point out that an algorithm which runs in worse than linear time (e.g., $O(n^2 (m + n))$) may be designed with weaker understanding of the separator structure. The upper bound (Theorem 5) can be obtained in other ways also, though we haven’t seen it mentioned anywhere else.

The only fact about chordal graphs, the reader should be familiar with, for reading this paper, apart from the notion of PEO, which will be explained in the next subsection, is a well known characterisation due to Dirac [2,6].

**Lemma 1.** A graph $G$ is a chordal graph if and only if every minimal separator of $G$ induces a complete subgraph.

### 1.1 PEO, Some Definitions, and Notations

Let $G = (V, E)$ be a connected undirected graph. Throughout this paper we will use $V$ for the set of vertices of $G$ and $E$ for the set of edges. $|V|$ and $|E|$ will be denoted by $n$ and $m$ respectively. $N(v)$ will denote the set of neighbours of $v$,
that is \( N(v) = \{ u \in V : (u, v) \in E \} \). For \( A \subseteq V \), we use \( N(A) \) to denote the set \( \bigcup_{v \in A} N(v) - A \). The subgraph of \( G \) induced by the nodes in \( A \) will be denoted by \( G(A) \).

A bijection \( f : V \to \{1, 2, \ldots, n\} \) is called an ordering of the vertices of \( G \). Then \( f(v) \) is referred to as the number associated with the vertex \( v \), or simply the number of \( v \) with respect to the ordering \( f \). Given an ordering \( f \) of a graph \( G \), we define the following terms.

**Definition 1.** Let \( A \subseteq V \). The highest(\( A \)) is defined to be the vertex with the highest number in \( A \). Similarly lowest(\( A \)) is the vertex in \( A \) with the lowest number.

**Definition 2.** A path \( P = (w_1, w_2, \ldots, w_k) \) in \( G \) is called an increasing path, iff \( f(w_1) < f(w_2) < \cdots < f(w_k) \). It is called a decreasing path iff \( f(w_1) > f(w_2) > \cdots > f(w_k) \). A single node can be considered as either increasing or decreasing.

**Definition 3.** A vertex \( u \in N(v) \) is called a higher neighbour of \( v \) iff \( f(u) > f(v) \). The set of higher neighbours of \( v \) will be denoted by \( N_h(v) \), ie

\[
N_h(v) = \{ u \in N(v) : f(u) > f(v) \}
\]

Similarly the set of lower neighbours of \( v \) is denoted by \( N_l(v) \).

\[
N_l(v) = \{ u \in N(v) : f(u) < f(v) \}
\]

\[ d_h(v) = |N_h(v)| \text{ and } d_l(v) = |N_l(v)|. \]

**Definition 4.** An ordering \( f \) of \( G \) is called a perfect elimination ordering (PEO) iff for each \( v \in V \), \( G(\{ v \} \cup N_h(v)) \) is a complete subgraph (clique) of \( G \). Then \( f(v) \) will be called the PEO number of \( v \).

Note that every graph may not have a PEO. In fact it is known that a graph \( G \) is chordal if and only if there exists a PEO for \( G \) \cite{6}. A \( O(m + n) \) algorithm to construct a PEO for a chordal graph \( G \) is given in \cite{21}. Note that there can be more than one PEO for a given chordal graph. The observations presented in this paper are valid with respect to any PEO. Therefore we just use \( PEO(v) \) to denote the PEO number of \( v \). It will be clear from the context which PEO is being referred to.

From now on we are only interested in chordal graphs. We will be using \( G \) to denote an arbitrary connected chordal graph. We close this subsection with one more definition.

**Definition 5.** A vertex \( v \) is called an interior vertex of \( G \), iff there exists a vertex \( u \notin N(v) \) with \( PEO(u) > PEO(v) \). Note that \( v \) is an interior vertex if and only if \( d_h(v) < n - PEO(v) \). We denote the set of interior vertices by \( V' \). That is, \( V' = \{ v \in V : d_h(v) < n - PEO(v) \} \).
2 The Minimum Separators

A chordless path from \( u \) to \( v \) is defined to be a path from \( u \) to \( v \) in \( G \) such that no two non consecutive nodes of the path are adjacent. The reader can easily verify that if there is a path between \( u \) and \( v \) then there is a chordless path also. For example, the shortest path between \( u \) and \( v \) has to be a chordless path.

Lemma 2. Let \( P = (w_1, w_2, \ldots, w_k) \) be a chordless path in a chordal graph \( G \) and let \( w_i = \text{highest}(P) \). Then \((w_1, w_2, \ldots, w_i)\) is an increasing path while \((w_i, w_{i+1}, \ldots, w_k)\) is a decreasing path.

Proof: Consider first the path \((w_1, w_2, \ldots, w_i)\). Since \( w_i = \text{highest}(P) \), this cannot be a decreasing path. If \( i \leq 2 \), obviously this is an increasing path. Let \( i > 2 \). If \((w_1, w_2, \ldots, w_i)\) is not an increasing path, then there should be a node \( w_j \in P, 1 < j < i \), such that \( \text{PEO}(w_j) < \text{PEO}(w_{j+1}) \) and \( \text{PEO}(w_j) < \text{PEO}(w_{j-1}) \). Then \((w_{j-1}, w_{j+1}) \subseteq N_h(w_j) \). By definition of PEO, \( N_h(w_j) \) forms a clique, and we have \((w_{j-1}, w_{j+1}) \subseteq E \). So \((w_{j-1}, w_{j+1})\) forms a chord for \( P \), contradicting the chordless property of \( P \). We infer that \((w_1, w_2, \ldots, w_i)\) is an increasing path. A similar argument shows that \((w_i, w_{i+1}, \ldots, w_k)\) should be a decreasing path. \(\square\)

Corollary 1. Let \( P=(w_1, w_2, \ldots, w_k) \) be a chordless path and \( w_k=\text{highest}(P) \). Then \( P \) is an increasing path.

Corollary 2. Let \( P = (w_1, w_2, \ldots, w_k) \) be a chordless path and \( \text{PEO}(w_1) < \text{PEO}(w_k) \). Then \( \text{PEO}(w_2) > \text{PEO}(w_1) \).

Lemma 3. Let \( P = (w_1, w_2, \ldots, w_k) \) be an increasing path in a chordal graph. Let \( \text{PEO}(u) > \text{PEO}(w_k) \) where \( u \in N(w_1) \). Then \( u \in N_h(w_k) \).

Proof: We prove the lemma by induction on \(|P|\), the number of nodes in \( P \). For \(|P| = 1\), the lemma is trivial. Assume that for all increasing paths with \(|P| = k-1\), where \( k > 1 \), the lemma is true. Let \( P = (w_1, w_2, \ldots, w_k) \) be an increasing path with \(|P| = k\). Note that we have \( \text{PEO}(u) > \text{PEO}(w_k) > \text{PEO}(w_{k-1}) \) since \( P \) is an increasing path. Applying the induction assumption on the \((k-1)-\)node path \((w_1, w_2, \ldots, w_{k-1})\), we get \( u \in N_h(w_{k-1}) \). Now \( \{w_k, u\} \subseteq N_h(w_{k-1}) \) and by definition of PEO, \( N_h(w_{k-1}) \) is a clique, and \((u, w_k) \in E \). Remembering \( \text{PEO}(u) > \text{PEO}(w_k) \), we conclude \( u \in N_h(w_k) \). \(\square\)

Lemma 4. Let \( A \subset V \) such that \( G(A) \) is connected. Let \( x = \text{highest}(A) \) and \( z = \text{lowest}(N(A)) \). Then if \( \text{PEO}(x) < \text{PEO}(z) \), \( N(A) = N_h(x) \).

Proof: First note that since \( x = \text{highest}(A) \), \( N_h(x) \cap A = \phi \) and therefore, \( N_h(x) \subseteq N(A) \). Now we will prove \( N(A) \subseteq N_h(x) \), from which we can conclude \( N(A) = N_h(x) \). Let \( y \in N(A) \). Then there is a \( w \in A \) such that \( y \in N(w) \). Also \( \text{PEO}(y) \geq \text{PEO}(z) > \text{PEO}(x) \). Consider a chordless path \( P = (w, \ldots, x) \), which is completely in \( G(A) \). Such a path exists, since \( G(A) \) is connected. Also since \( x = \text{highest}(P) \), by corollary 4, \( P \) is an increasing path. Then by lemma 3, \( y \in N_h(x) \). Thus we conclude that \( N(A) \subseteq N_h(x) \). It follows that \( N(A) = N_h(x) \). \(\square\)
Lemma 5. Let $x$ be an interior vertex in a chordal graph $G$, ie $d_h(x) < n - PEO(x)$. Then $N_h(x)$ is a separator of $G$. In particular, $N_h(x)$ is a $(x - y)$ separator for all $y \notin N_h(x)$ with $PEO(y) > PEO(x)$.

Proof: Let $PEO(x) = t$. Consider the induced subgraph $G(V_t)$ on the set of nodes $V_t = \{u \in V : PEO(u) \leq t\}$. Let $C_i$ be the connected component of $G(V_t)$ which contains $x$. Obviously $x = highest(C_i)$. Since $N(C_i) \subseteq V - V_t$, clearly every node $z \in N(C_i)$ has $PEO(z) > t = PEO(x)$. Now by lemma 4 $N(C_i) = N_h(x)$. Since $d_h(x) < n - PEO(x)$, there should be a node $y$ such that $PEO(y) > PEO(x)$ and $y \notin N_h(x)$. Clearly $y \notin C_i$ since $x = highest(C_i)$. Therefore $N_h(x)$ separates $x$ from $y$, from which the lemma follows. $\square$

Lemma 6. Let $\kappa(G)$ be the connectivity of a chordal graph $G$. Then the number of interior nodes, $|V'|$ is at most $n - \kappa(G) - 1$.

Proof: Let $x$ be an interior node of $G$. Note that by lemma 5, $N_h(x)$ is a $(x - y)$ separator for some $y$ with $PEO(y) > PEO(x)$. Then if $PEO(x) \geq n - \kappa(G)$, $|N_h(x)| < \kappa(G)$, which contradicts the definition of $\kappa(G)$. Thus we infer that $PEO(x) \leq n - \kappa(G) - 1$. It follows that $|V'| \leq n - \kappa(G) - 1$. $\square$

Theorem 1. (Characterisation of minimum separators) Let $G$ be a chordal graph with $\kappa(G) = k$. Let $M$ be the set of all minimum sized separators of $G$. Then

$$M = \{S \subseteq V : S = N_h(x) \text{ for some } x \in V' \text{ such that } d_h(x) = k\}$$

where $V' = \{u \in V : d_h(u) < n - PEO(u)\}$, the set of interior nodes of $G$.

Proof: First we prove that every $S \in M$ is indeed a minimum separator of $G$. By lemma 5, $S$ is a separator of $G$. Since we have $k = d_h(x) = |N_h(x)| = |S|$, $S$ is a minimum separator, since $\kappa(G) = k$.

Now let $S$ be a minimum separator, ie $|S| = k$. Also let $V$ be partitioned into $A, S$ and $B$ by $S$. Let $x = highest(A)$ and let $y = highest(B)$. Without loss of generality let $PEO(x) < PEO(y)$. Then $x$ is an interior vertex of $G$, since $y \notin N(x)$ and $PEO(y) > PEO(x)$. By lemma 5 $N_h(x)$ is a separator. Also since $\kappa(G) = k$, $|N_h(x)| \geq k$. Clearly $N_h(x) \subseteq S$, since $x = highest(A)$. Then, since $|S| = k$, we conclude that $N_h(x) = S$, and $d_h(x) = k$. Thus $S \in M$. $\square$.

Corollary 3. Let $G$ be a chordal graph, which is not a clique. Then $\kappa(G) = \min_{u \in V'} d_h(u)$, where $V' = \{u \in V : d_h(u) < n - PEO(u)\}$, the set of interior nodes.

Algorithm to enumerate minimum separators

Input : A connected chordal graph $G$, which is not complete.
Step 1: Compute a PEO of $G$.
Step 2: Find $d_h(u)$ for each $u \in V$.
Step 3: Let $V' = \{u : d_h(u) < n - PEO(u)\}$.
Step 4: $\kappa(G) = \min_{u \in V'} d_h(u)$. 
Step 5: For each \( u \in V' \), if \( d_h(u) = \kappa(G) \) output \( N_h(u) \).

It can easily be verified that the above algorithm runs in \( O(m + n) \) time. Computing the PEO takes \( O(m+n) \) time [21]. Steps 2 and 5 can be implemented in \( O(m) \) time, since it only involves traversing the neighbour lists of each node. Steps 3 and 4 takes \( O(n) \) time.

The correctness of step 4 is ensured by corollary 3 and the correctness of step 5, namely the sets output are exactly the minimum separators is ensured by Theorem 1.

3 The Minimal Separators

Theorem 2. (Characterisation of minimal separators) \( G \) is a minimal separator of a chordal graph \( G \) if and only if there exist two nonadjacent vertices \( a, b \in V \), such that \( \text{PEO}(a) < \text{PEO}(b) \), and \( S = N_h(a) = N_h(a) \cap N(b) \).

Proof: First assume that there are two nonadjacent nodes \( a, b \) with \( \text{PEO}(a) < \text{PEO}(b) \), such that \( S = N_h(a) = N_h(a) \cap N(b) \). Clearly \( a \) is an interior node and \( S = N_h(a) \) is a separator which separates \( a \) from \( b \), by lemma 5. Clearly \( S \) is minimal since \( S = N_h(a) \cap N(b) \) : removal of any proper subset of \( S \) leaves out a path from \( a \) to \( b \).

Next, let \( S \) be a minimal separator which separates \( v_1 \) from \( v_2 \). Let \( A \) and \( B \) be the connected components of \( G(V - S) \), which contains \( v_1 \) and \( v_2 \) respectively. First we show that there exists a node \( a \in A \) and a node \( b \in B \) such that \( S = N(a) \cap N(b) \).

First note that \( S = N(A) = N(B) \). To see this, observe that because of the minimality of \( S, S \subseteq N(A) \). Clearly since \( A \) is a connected component of \( G(V - S) \), \( N(A) \subseteq S \), also. Thus we have \( S = N(A) \). By similar arguments \( S = N(B) \).

Let \( z = \text{lowest}(S) \), and \( x = \text{highest}(A) \). If \( \text{PEO}(x) < \text{PEO}(z) \), then since \( N(A) = S \), by lemma 4 we have \( S = N(A) = N_h(x) \subseteq N(x) \). In this case let \( a = x \).

If \( \text{PEO}(x) > \text{PEO}(z) \), consider a chordless path from \( z \) to \( x \), which is completely in the induced subgraph \( G(A \cup \{z\}) \). (Observe that such a path exists, because \( G(A \cup \{z\}) \) is connected since \( z \in N(A) \)). Let \( u \) be the second node in this chordless path starting from \( z \). By corollary 2, \( \text{PEO}(u) > \text{PEO}(z) \). Also by lemma 4 \( S \) forms a complete subgraph of \( G \). Thus we have \( S - \{z\} \cup \{u\} \subseteq N_h(z) \), since \( z = \text{lowest}(S) \). By definition of PEO, \( N_h(z) \cup \{z\} \) form a complete subgraph and thus we have \( S \subseteq N(u) \). In this case let \( a = u \).

By similar arguments as above, we get a node \( b \in B \) such that \( S \subseteq N(b) \). Now we show that \( S = N(a) \cap N(b) \). Clearly \( S \subseteq N(a) \cap N(b) \). Suppose there is a node \( v \in (N(a) \cap N(b) - S) \). Then \( S \) can not separate \( a \) from \( b \), which is a contradiction. Thus we have \( S = N(a) \cap N(b) \).

Without loss of generality, assume that \( \text{PEO}(a) < \text{PEO}(b) \). We will now show that \( S = N_h(a) \cap N(b) \). First we claim that \( \text{PEO}(a) < \text{PEO}(z) \). Otherwise let \( \text{PEO}(b) > \text{PEO}(a) > \text{PEO}(z) \). Note that since \( a \) and \( b \) are neighbours of every node in \( S \), we have \( a, b \in N_h(z) \). By definition of PEO, there will be an
edge between \( a \) and \( b \) which is a contradiction since \( S \) separates \( a \) from \( b \). Thus \( PEO(a) < PEO(z) \), and we have \( S \subseteq N_h(a) \). It follows that \( S = N_h(a) \cap N(b) \).

Finally, since we have established \( PEO(a) < PEO(z) \), it follows that \( a = x = \text{highest}(A) \). (Remember we chose \( a \), from two choices: one was \( x \), the \( \text{highest}(A) \) and the other was \( u \), the second node in the chordless path from \( z \) to \( x \). In the latter case, \( PEO(a) > PEO(z) \).) Thus we have \( S = N_h(a) \).

\[ \square \]

The above theorem says that every minimal separator has to be \( N_h(u) \) for some interior vertex. Our algorithm to enumerate all minimal separators examines \( N_h(u) \) for each interior vertex \( u \), and decides whether it is a minimal separator or not. We consider a node \( w \) to be a witness for \( N_h(u) \), if the existence of \( w \) proves that \( N_h(u) \) is a minimal separator. For example, if we can find a node \( w \), nonadjacent to \( u \) with \( PEO(w) > PEO(u) \) and \( N(w) \supseteq N_h(u) \), then by Theorem 2, \( w \) is a witness for \( N_h(u) \). Thus the issue in designing the algorithm is to efficiently identify a witness, if one exists.

In our algorithm, we make use of only two types of witnesses. These types are defined in terms of a special node in \( N_h(u) \), namely \( z = \text{lowest}(N_h(u)) \).

**Definition 6.** First type witness: A node \( w \) is defined to be a first type witness for \( N_h(u) \), iff \( w \in N_h(z) \setminus N(u) \).

**Lemma 7. (First type witnesses are indeed witnesses)** If there exists a node \( w \in N_h(z) \setminus N(u) \), then \( N_h(u) \) is a minimal separator.

**Proof:** First note that \( w \in N_h(z) \setminus N(u) \) is not adjacent to \( u \) and since \( w \in N_h(z) \), \( PEO(w) > PEO(u) \). Now note that since \( N_h(u) \) is a clique, and \( z = \text{lowest}(N_h(u)) \), we have \( N_h(u) \setminus \{ z \} \subseteq N_h(z) \). Also \( w \in N_h(z) \). Since by definition of \( PEO \), \( \{ z \} \cup N_h(z) \) forms a clique, it follows that \( N_h(u) \subseteq N(w) \). Then by Theorem 2, \( w \) is a witness for \( N_h(u) \).

**Definition 7.** Second type witness: A node \( w \neq u \), is defined to be a second type witness iff \( w \in N_l(z) \setminus N(u) \), and \( N_h(w) = N_h(u) \).

**Lemma 8. (Second type witnesses are indeed witnesses)** If there exists \( w \neq u \), \( w \in N_l(z) \setminus N(u) \), such that \( N_h(w) = N_h(u) \), then \( N_h(u) \) is a minimal separator.

**Proof:** First note that \( w \) is nonadjacent to \( u \). If \( PEO(w) > PEO(u) \), clearly since \( N_h(u) \cap N(w) = N_h(w) = N_h(u) \), by Theorem 2, \( N_h(u) \) is a minimal separator and \( w \) is a witness for \( N_h(u) \). If \( PEO(w) < PEO(u) \), we just have to interchange the roles of \( u \) and \( w \), and we again get \( N_h(u) = N_h(w) \) is a minimal separator. \( \square \)

The first and second type witnesses certainly do not exhaust the set of all possible witnesses. But it turns out that these two types are sufficient for our algorithm. This is because of the following lemma, which assures that, if first type witnesses are not available, then a second type witness is guaranteed to exist, provided the \( N_h(u) \) in question is a minimal separator.
Lemma 9. Let $N_h(u)$ be a minimal separator. If $N_h(z) - N(u) = \phi$, then there exists a node $w \neq u$, such that $w \in N_l(z) - N(u)$, and $N_h(w) = N_h(u)$.

Proof: Since $N_h(u)$ is a minimal separator, there exists a node $v \in V$, such that $N_h(u)$ minimally separates $v$ from $u$. (If $N_h(u)$ is a minimal $(v_1 - v_2)$ separator, at least one of them, either $v_1$ or $v_2$, should be in a different connected component than that of $u$, when $N_h(u)$ is removed from $G$. Then clearly $N_h(u)$ minimally separates $v$ from that node.) Let $A$ be the connected component of $G(V - N_h(u))$ which contains $v$. Clearly $u \not\in A$. First note that by minimality of the separator $N_h(u)$, $N_h(u) = N(A)$. Let $x = \text{highest}(A)$. We claim that $\text{PEO}(x) < \text{PEO}(z)$. Otherwise, if $\text{PEO}(x) > \text{PEO}(z)$, consider a chordless path $P = (z, w_1, \ldots, x)$ in $G(A \cup \{z\})$, which is guaranteed to exist since $G(A \cup \{z\})$ is connected. Then clearly $w_1 \in N_h(z)$, by corollary 2. Also $w_1 \not\in N(u)$, since $w_1 \in A$ and $u$ is separated from $A$ by $N_h(u)$. Thus $w_1 \in N_h(z) - N(u)$, contradicting the assumption that $N_h(z) - N(u) = \phi$. We infer that $\text{PEO}(x) < \text{PEO}(z)$. We conclude from lemma 4 that $N_h(x) = N(A) = N_h(u)$. Also since $x \in N_l(z)$ and $x \not\in N(u)$, we have $x \in N_l(z) - N(u)$. The lemma follows. □

Finally, the reason why we consider these special kind of witnesses, namely the second type witnesses, is exactly that, they are easy to identify. The following lemma explains this.

Lemma 10. Suppose that $N_h(z) - N(u) = \phi$. A node $w \neq u$, $w \in N_l(z) - N(u)$ is a second type witness if and only if $d_h(w) = d_h(u)$ and $\text{lowest}(N_h(w)) = z$.

Proof: If $w$ is a second type witness, ie, if $N_h(w) = N_h(u)$, clearly we have $d_h(u) = d_h(w)$ and $\text{lowest}(N_h(w)) = z$. Now suppose that there is a node $w \in N_l(z) - N(u)$, with $d_h(u) = d_h(w)$ and $\text{lowest}(N_h(w)) = z$. We first show that $N_h(w) \subseteq N_h(u)$. Let $y \in N_h(w)$. Noting that by the assumption $z = \text{lowest}(N_h(w))$, $z \in N_h(w)$ also, we get $(z, y) \in E$, since $N_h(w)$ should form a clique. But $z = \text{lowest}(N_h(w))$, therefore $\text{PEO}(y) > \text{PEO}(z)$, ie $y \in N_h(z)$. Now if $y \not\in N_h(u)$, then $y \in N_h(z) - N_h(u) = N_h(z) - N(u)$, which contradicts the assumption that $N_h(z) - N(u) = \phi$. Thus $y \in N_h(u)$. It follows that $N_h(w) \subseteq N_h(u)$. But since $d_h(u) = d_h(w)$, it has to be the case that $N_h(u) = N_h(w)$. That is, $w$ is a second type witness of $N_h(u)$. □

Algorithm To Enumerate Minimal Separators

Step 1: Find a PEO of $G$.

Step 2: (Preprocessing:) For each node $u \in V$, find $d_h(u)$ and $z_u = \text{lowest}(N_h(u))$. Store these information with the node, for future use.

Step 3: For each $u \in V'$ with nonempty $N_h(u)$, do:

Step 3.1: (Looking for first type witness:) If $N_h(z_u) - N(u) = \phi$, output $N_h(u)$.

Step 3.2: (Looking for second type witness:) If $N_h(z_u) - N(u) = \phi$, then if there is a node $w \neq u$ and $w \in N_l(z_u) - N(u)$, such that $d_h(u) = d_h(w)$ and $z_w = z_u$, then output $N_h(u)$.

Theorem 3. The above algorithm outputs exactly the set of all minimal separators.
Proof: First note that the algorithm outputs $N_h(u)$ only if a witness is found. Therefore it outputs only minimal separators. Now every minimal separator is $N_h(u)$ for some $u \in V'$, by Theorem [2] and has either a first type witness or second type witness by lemma [9]. Therefore every minimal separator will be output. □

**Theorem 4.** The above algorithm runs in $O(m + n)$ time.

Proof: Note that implementing step 2 and 3 only involves traversing the list of neighbours of each node a constant number of times. Since $\Sigma_{u \in V} d(u) = 2m$, this can be done in $O(m)$ time. Step 1 can be implemented in $O(m + n)$ time. □

### 4 Upper Bound for Number of Minimal Separators of Chordal Graphs

**Theorem 5.** Number of minimum separators $\leq n - \kappa(G) - 1$.

Proof: Since every minimum separator is a minimal separator, we just have to show the second inequality. It is clear from Theorem [2] that each minimal separator is $N_h(u)$ for some interior vertex $u$. But the number of interior vertices is $\leq n - \kappa(G) - 1$, by lemma [6]. Hence the result. □

Finally we show that the above inequality can be tight, by constructing a graph $G_n(V_n, E_n)$, with $\kappa(G) = k$, and $n \geq k + 2$, for $k \geq 1$, such that the number of minimum separators $= n - k - 1$. This clearly proves that for $G_n$, the number of minimum separators $= n - \kappa(G) - 1$. For this we just have to show that $V'_n$, the set of interior nodes of $G_n$, contain exactly $n - k - 1$ nodes, and for each $u \in V'_n$, $d_h(u) = k$ (see Theorem [1]).

First let $l = k + 2$. Let $V_l = \{v_1, v_2, \ldots, v_l\}$. Define $G_l(V_l, E_l)$ to be the graph obtained by removing the edge $(v_1, v_l)$ from the complete graph on $V_l$. Consider an ordering $f_i$ of $G_l$ defined by $f_i(v_i) = i$, for $1 \leq i \leq l$. We claim that $G_l$ is a chordal graph, and $f_i$ is a PEO. We leave the details to the reader. Also it is easy to verify that $\kappa(G) = k$ and $V'_l = \{v_1\}$. Thus we have just one, minimum separator for $G_l$.

Now we inductively construct $G_n(V_n, E_n)$. Let $G_{n-1}(V_{n-1}, E_{n-1})$ be a chordal graph with PEO defined by $f_{n-1}(v_i) = i$, for $1 \leq i \leq n - 1$. Let $\kappa(G_{n-1}) = k$ and let $V'_{n-1} = \{v_1, v_2, \ldots, v_{n-k-2}\}$. Also let for all $u \in V'_{n-1}$, $d_h(u) = k$ and $N_h(u)$ distinct from one another. In other words, the number of minimum separators of $G_{n-1}$ is exactly $(n - 1) - k - 1$.

Now let $V_n = V_{n-1} \cup \{v_n\}$, where $v_n$ is a new node. Let $f_n(v_i) = i$, for $1 \leq i \leq n$ and $E_n = E_{n-1} \cup \{(v_i, v_n) : v_i \in V_{n-1} \text{ with } f_n(v_i) = i > n - k - 1\}$. We claim that $G_n$ is a chordal graph with a PEO defined by $f_n$. Also it is not difficult to show that $\kappa(G_n) = k$. We leave the details to the reader.

Clearly $V'_n = V'_{n-1} \cup \{v_{n-k-1}\}$. Note that $d_h(v_{n-k-1}) = k$. Thus we have $|V'_n| = n - k - 1$, and $\forall v \in V'_n$, $d_h(v) = \kappa(G)$ and $N_h(v_{n-k-1}) \neq N_h(u) \forall u \in V'_{n-1}$. Thus $G_n$ has exactly $n - \kappa(G) - 1$ minimum separators by Theorem [1]
References

Graph Separators: A Parameterized View

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Abstract. Graph separation is a well-known tool to make (hard) graph problems accessible for a divide and conquer approach. We show how to use graph separator theorems in order to develop fixed parameter algorithms for many well-known NP-hard (planar) graph problems. We coin the key notion of glueable select&verify graph problems and derive from that a prospective way to easily check whether a planar graph problem will allow for a fixed parameter divide and conquer algorithm of running time $c^{\sqrt{k}} \cdot n^{O(1)}$ for a constant $c$.

1 Introduction

Algorithm designers are often faced with problems which, when viewed from classical computational complexity theory, are “intractable.” More formally speaking, these problems can be shown to be NP-hard. In many applications, however, a certain part (called the parameter) of the whole problem can be identified which tends to be of small size $k$ when compared with the size $n$ of the whole problem instance. This leads to the study of parameterized complexity [7].

Fixed parameter tractability. Formally, one terms a (parameterized) problem fixed parameter tractable if it allows for a solving algorithm running in time $f(k)n^{O(1)}$ on input instance $(I,k)$, where $n = |I|$ and $f$ is an arbitrary function only depending on $k$. We will also term such algorithms “$f(k)$-algorithms” for brevity, focusing on the exponential part of the running time bound. The associated complexity class is called FPT. Of course, designing fixed parameter algorithms with a “small” function $f$ is desirable. To our knowledge, so far, only one non-trivial fixed parameter tractability result where the corresponding function $f$ is sublinear in the exponent, namely $f(k) = c^{\sqrt{k}}$ [11], is known: PLANAR DOMINATING SET. Similar results hold for closely related problems such as FACE COVER, PLANAR INDEPENDENT DOMINATING SET, PLANAR WEIGHTED DOMINATING SET, etc. [11]. In the companion paper [2], we proved similar results for a much broader class of planar graph problems, presenting a general methodology based on concepts such as tree decompositions and bounded outerplanarity.

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Here, we will also discuss a rather general approach for obtaining parameterized graph algorithms running in time $\mathcal{O}(c^k \cdot q(n))$ for sublinear functions $e$, i.e., $e(k) \in o(k)$. More precisely, we investigate the use of (planar) separator theorems in this context, yielding an alternative and conceptually rather different framework in comparison with [2].

**General outline.** It is a crucial goal throughout the paper not to narrowly stick to problem-specific approaches, but to try to widen the techniques as far as possible. More specifically, we show how to use separator theorems for different graph classes, such as, e.g., the well-known planar separator theorem due to Lipton and Tarjan [8], in combination with known algorithms for obtaining (linear size) problem kernels in order to obtain fixed parameter divide and conquer algorithms. Special care is taken for the dependency of the running time on the “graph separator parameters” (and how they influence the recurrences in the running time analysis). We consider a broad class of problems that can be attacked by this approach, namely, in principle, all so-called glueable select&verify problems such as, e.g., **planar dominating set**. Also, we exhibit the influence on the running time analysis of so-called cycle separators. Although the constants achieved in our setting so far seem to be too large in order to yield practical algorithms, our approach provides a general, sound, mathematical formalization of a rich class of problems that allow for divide and conquer fixed parameter algorithms. Our methodology seems to leave much room for improvement in many directions. For instance, we introduce the novel concept of problem cores that can replace problem kernels in our setting. Finally, we give a push to the study of subclasses of the parameterized complexity class FPT. In this sense, our work also might serve as a starting point for more algorithmic (also concerning graph theory with respect to separator theorems), as well as more structural complexity-theoretic lines of future research in parameterized complexity.

Due to the lack of space, several details are deferred to the long version.

## 2 Basic Definitions and Preliminaries

We consider undirected graphs $G = (V,E)$, $V$ denoting the vertex set and $E$ denoting the edge set. We only consider simple graphs (i.e., with no double edges) without self-loops. Sometimes, we refer to $V$ by $V(G)$ in order to emphasize that $V$ is the vertex set of graph $G$; by $N(v)$ we refer to the set of vertices adjacent to $v$. $G[D]$ denotes the subgraph induced by a vertex set $D$. For graphs $G_1 = (V_1,E_1)$ and $G_2 = (V_2,E_2)$, by $G_1 \cap G_2$, we denote the graph $(V_1 \cap V_2, E_1 \cap E_2)$. $G' = (V',E')$ is a subgraph of $G = (V,E)$, denoted by $G' \subseteq G$, if $V' \subseteq V$ and $E' \subseteq E$. In this paper, we only consider graph classes, denoted by $\mathcal{G}$, that are closed under taking subgraphs. The most important among these graph classes is that of **planar** graphs, i.e., graphs that have a drawing in the plane without

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1. Here and in the following, $n$ is the number of vertices of the input graph and $k$ is the parameter of the considered graph problem.
2. Lipton and Tarjan [9] only describe a solution for the structurally much simpler **planar independent set** in detail.
edge crossings. Among others, we study the following vertex subsets of a graph $G = (V, E)$: A *vertex cover* $C \subseteq V$ satisfies that every edge of $G$ has at least one endpoint in $C$. An *independent set* is a set of pairwise nonadjacent vertices. A *dominating set* $D \subseteq V$ obeys that each of the rest of the vertices in $G$ has at least one neighbor in $D$. The corresponding problems are denoted by (PLANAR) VERTEX COVER, INDEPENDENT SET, and DOMINATING SET.

Further conventions: we write $A + B$ to denote the disjoint union of sets $A$ and $B$. We let $0 \cdot (\pm \infty) = 0$.

**Linear problem kernels.** Let $\mathcal{L}$ be a parameterized problem, i.e., $\mathcal{L}$ is a subset of $\Sigma^* \times \mathbb{N}$

Reduction to problem kernel, then, means to replace instance $(I, k) \in \Sigma^* \times \mathbb{N}$ by a “reduced” instance $(I', k') \in \Sigma^* \times \mathbb{N}$ (which we call problem kernel) such that $k' \leq c \cdot k, |I'| \leq p(k)$ with constant $c$, some function $p$ only depending on $k$, and $(I, k) \in \mathcal{L}$ iff $(I', k') \in \mathcal{L}$. Furthermore, we require that the reduction from $(I, k)$ to $(I', k')$ is computable in polynomial time $T_K(|I|, k)$.

Often, the best one can hope for is that the problem kernel is size linear in $k$, a so-called *linear problem kernel*. For instance, using a theorem of Nemhauser and Trotter, Chen et al. \cite{3} observed a problem kernel of size $2k$ for VERTEX COVER on general (not necessarily planar) graphs. Furthermore, PLANAR INDEPENDENT SET has a problem kernel of size $4k$ due to the four color theorem. Once having a linear size problem kernel, it is fairly easy to use our framework to get $c^{\sqrt{k}}$-algorithms for these problems based upon the famous planar separator theorem \cite{8,9}. The constant factor in the problem kernel size directly influences the value of the exponential base. Hence, lowering the kernel size is a crucial goal.

**Classical separator theorems.** Let $G = (V, E)$ be an undirected graph. A separator $S \subseteq V$ of $G$ partitions $V$ into two parts $A_1 \subseteq V$ and $A_2 \subseteq V$ such that $A_1 + S + A_2 = V$, and no edge joins vertices in $A_1$ and $A_2$; $(A_1, S, A_2)$ is called a separation of $G$. When we restrict our attention to planar graphs, $S$ is called cycle separator if it forms a cycle in some triangulation of $G$. According to Lipton and Tarjan \cite{8}, an $f(\cdot)$-separator theorem (with constants $\alpha < 1$, $\beta > 0$) for a class $\mathcal{G}$ of graphs which is closed under taking subgraphs is a theorem of the following form: If $G$ is any $n$-vertex graph in $\mathcal{G}$, then there is a separation $(A_1, S, A_2)$ of $G$ such that neither $A_1$ nor $A_2$ contains more than $\alpha n$ vertices, and $S$ contains no more than $\beta f(n)$ vertices.

Stated in this framework, the planar separator theorem due to Lipton and Tarjan \cite{8} is a $\sqrt{\cdot}$-separator theorem with constants $\alpha = 2/3$ and $\beta = 2\sqrt{2}$. The current record for $\alpha = 2/3$ is $\beta \approx 1.97$ \cite{6}. Similar $\sqrt{\cdot}$-separator theorems are also known for other graph classes, e.g., for the class of graphs of bounded genus, see \cite{5}. It is also possible to incorporate weights in most separator theorems. We refer to them as $f(\cdot)$-separator theorems for weighted graphs.

\footnote{In this paper, we assume the parameter to be a positive integer although, in general, it might also be from an arbitrary language (e.g., being a subgraph).}
3 Glueable Graph Problems

Select&verify graph problems. A set $G$ of tuples $(G, k)$, $G$ an undirected graph with vertex set $V = \{v_1, \ldots, v_n\}$ and $k$ a nonnegative real number, is called a select&verify (graph) problem if there exists a pair $(P, \text{opt})$ with $\text{opt} \in \{\text{min}, \text{max}\}$, such that $P$ is a function that assigns to $G$ a polynomial time computable function of the form $P_G = P_G^{\text{sel}} + P_G^{\text{ver}}$, where $P_G^{\text{sel}} : \{0, 1\}^n \to \mathbb{R}_+$, $P_G^{\text{ver}} : \{0, 1\}^n \to \{0, \pm \infty\}$, and

$$(G, k) \in G \iff \begin{cases} \text{opt}_{x \in \{0, 1\}^n} P_G(x) \leq k & \text{if } \text{opt} = \text{min}, \\ \text{opt}_{x \in \{0, 1\}^n} P_G(x) \geq k & \text{if } \text{opt} = \text{max}. \end{cases}$$

For $x = (x_1, \ldots, x_n) \in \{0, 1\}^n$ with $P_G(x) \leq k$ if $\text{opt} = \text{min}$ and with $P_G(x) \geq k$ if $\text{opt} = \text{max}$, the vertex set selected by $x$ and verified by $P_G$ is $\{v_i \in V \mid x_i = 1, 1 \leq i \leq n\}$. A vector $x$ is called admissible if $P_G^{\text{ver}}(x) = 0$.

The intuition behind the term $P = P^{\text{sel}} + P^{\text{ver}}$ is that the “selecting function” $P^{\text{sel}}$ counts the size of the selected set of vertices and the “verifying function” $P^{\text{ver}}$ verifies whether this choice of vertices is an admissible solution. Every select&verify graph problem that additionally admits a problem kernel of size $p(k)$ is solvable in time $O(2^{p(k)}p(k) + T_K(n, k))$.

We now give two examples for select&verify problems by specifying the function $P_G = P_G^{\text{sel}} + P_G^{\text{ver}}$. In both cases the “selecting function” for a graph $G = (V, E)$ will be $P_G^{\text{sel}} = \sum_{v_i \in V} x_i$. Firstly, in the case of VERTEX COVER, we have $\text{opt} = \text{min}$ and choose $P_G^{\text{ver}}(x) = \sum_{\{v_i, v_j\} \in E} \infty \cdot (1 - x_i)(1 - x_j)$. Thus, $P_G(x) \leq k$ guarantees a size at most $k$ vertex cover set. Secondly, for DOMINATING SET, we have $P_G^{\text{ver}}(x) = \sum_{v_i \in V} (\infty \cdot (1 - x_i) \cdot \prod_{\{v_i, v_j\} \in E} (1 - x_j))$.

We will also need a notion of select&verify problems where the “selecting function” and the “verifying function” operate on a subgraph of the given graph: Let $P = P^{\text{sel}} + P^{\text{ver}}$ be the function of a select&verify problem. For an $n$-vertex graph $G$ and subgraphs $G^{\text{ver}} = (V^{\text{ver}}, E^{\text{ver}}), G^{\text{sel}} = (V^{\text{sel}}, E^{\text{sel}}) \subseteq G$, we let

$$P_G^{\text{ver}}(x \mid G^{\text{sel}}) := P_G^{\text{ver}}(\pi_{V^{\text{ver}}}(x)) + P_G^{\text{sel}}(\pi_{V^{\text{sel}}}(x)),$$

where $\pi_{V'}$ is the projection of the vector $x \in \{0, 1\}^n$ to the variables corresponding to the vertices in $V' \subseteq V$.

Glueability. We are going to solve graph problems, slicing the given graph into small pieces with the help of small separators. The separators will serve as boundaries between the different graph parts into which the graph is split. For each possible assignment of the vertices in the separators, we want to – independently – solve the corresponding problems on the graph parts and then reconstruct a solution for the whole graph by “gluing” together the solutions for the graph parts. We need to assign colors to the separator vertices in the course of the algorithm. Hence, our algorithm has to be designed in such a manner that it can also cope with colored graphs. In general (e.g., in the case of the DOMINATING SET problem), it is not sufficient to simply use the two colors 1 (for encoding “in the selected set”) and 0 (for “not in the selected set”).
Let us introduce some auxiliary notions. Let $G = (V, E)$ be an undirected graph and let $C_0, C_1$ be finite, disjoint sets. A $C_0$-$C_1$-coloring of $G$ is a function $\chi : V \rightarrow C_0 + C_1 + \{\#\}$ for $V' \subseteq V$, a function $\chi : V' \rightarrow C_0 + C_1$ can naturally be extended to a $C_0$-$C_1$-coloring of $G$ by setting $\chi(v) = \#$ for all $v \in V \setminus V'$.

Consider a vector $x \in \{0, 1\}^{|V|}$. Let $\chi$ be a $C_0$-$C_1$-coloring of $G$. Then, $x$ is consistent with $\chi$ if, for $i = 0, 1$ and $j = 1, \ldots, |V|$, $\chi(v_j) \in C_i \Rightarrow x_j = i$.

If $\chi$ is a $C_0$-$C_1$-coloring of $G$ and if $\chi'$ is a $C_0'$-$C_1'$-coloring of $G$, then $\chi$ is preserved by $\chi'$, written $\chi \leadsto \chi'$, if $\forall v \in V \forall i \in \{0, 1\} (\chi(v) \in C_i \Rightarrow \chi'(v) \in C'_i)$.

In the next section, when doing the divide and conquer approach with a given separator, we will deal with colorings on two different color sets: one color set $C^{\text{int}} := C_0^{\text{int}} + C_1^{\text{int}} + \{\#\}$ of internal colors that will be used for the assignments of colors to the separator vertices and a color set $C^{\text{ext}} := C_0^{\text{ext}} + C_1^{\text{ext}} + \{\#\}$ of external colors that will be used for handing down the information in the divide-step of the algorithm. The idea is that, in each recursive step, we will be confronted with a graph “pre-colored” with external colors. Every function $\oplus$ that assigns to a pair $(\chi^{\text{ext}}, \chi^{\text{int}})$ with $\chi^{\text{ext}} : V \rightarrow C^{\text{ext}}, \chi^{\text{int}} : V \rightarrow C^{\text{int}}, \chi^{\text{ext}} \leadsto \chi^{\text{int}}$, a $(C_0^{\text{ext}}, C_1^{\text{ext}})$-coloring $\chi^{\text{ext}} \oplus \chi^{\text{int}}$ is called a recoloring if $\chi^{\text{int}} \leadsto \chi^{\text{ext}} \oplus \chi^{\text{int}}$.

From the point of view of recursion, $\chi^{\text{ext}}$ is the pre-coloring which a certain recursion instance “receives” from the calling instance and $\chi^{\text{int}}$ represents a coloring which this instance assigns to a certain part of the graph. The coloring $\chi^{\text{ext}} \oplus \chi^{\text{int}}$ is handed down in the recursion.

We now introduce the central notion of “glueable” select&verify problems. This formalizes those problems that can be solved with separator based divide and conquer techniques as described above.

**Definition 1.** A select&verify problem $\mathcal{G}$ given by $(P, \text{opt})$ is glueable with $\sigma$ colors if there exist

- a color set $C^{\text{int}} := C_0^{\text{int}} + C_1^{\text{int}} + \{\#\}$ of internal colors with $|C_0^{\text{int}} + C_1^{\text{int}}| = \sigma$;
- a color set $C^{\text{ext}} := C_0^{\text{ext}} + C_1^{\text{ext}} + \{\#\}$ of external colors;
- a polynomial time computable function $h : (\mathbb{R}_+ \cup \{\pm \infty\})^3 \rightarrow \mathbb{R}_+ \cup \{\pm \infty\}$;

and, for every $n$-vertex graph $G = (V, E)$ and subgraphs $G^{\text{ver}}, G^{\text{sel}} \subseteq G$ with a separation $(A_1, S, A_2)$ of $G^{\text{ver}}$, we find

- recolorings $\oplus_X$ for each $X \in \{A_1, S, A_2\}$, and
- for each internal coloring $\chi^{\text{int}} : S \rightarrow C^{\text{int}}$, subgraphs $G^{\text{ver}}(\chi^{\text{int}})$ of $G^{\text{ver}}$ with $G^{\text{ver}}[A_i] \subseteq G^{\text{ver}}(\chi^{\text{int}}) \subseteq G^{\text{ver}}[A_i + S]$ for $i = 1, 2$, and subgraphs $G^{\text{ver}}(\chi^{\text{int}})$ of $G^{\text{ver}}$ with $G^{\text{ver}}(\chi^{\text{int}}) \subseteq G^{\text{ver}}[S]$

such that, for each external coloring $\chi^{\text{ext}} : V \rightarrow C^{\text{ext}}$,

\[
\text{opt}\{P_{G^{\text{ver}}}(x \mid G^{\text{sel}}) \mid x \in \{0, 1\}^n \wedge x \sim \chi^{\text{ext}}\} = \text{opt}_{\chi^{\text{ext}} : S \rightarrow C_0^{\text{int}} + C_1^{\text{int}}} h(\text{Eval}_{A_1}(\chi^{\text{int}}), \text{Eval}_{S}(\chi^{\text{int}}), \text{Eval}_{A_2}(\chi^{\text{int}})).
\]

\footnote{The symbol $\#$ will be used for the undefined (i.e., not yet defined) color.}
Here, $\text{Eval}_X(\cdot)$ for $X \in \{A_1, S, A_2\}$ is of the form $\text{Eval}_X(\chi^{\text{int}}) = \text{opt}\{P(G^{\text{ver}}(\chi^{\text{int}}))(x) \mid G^{\text{ver}}[X] \cap G^{\text{sel}}) \mid x \in \{0, 1\}^n \land x \sim (\chi^{\text{ext}} \oplus x \chi^{\text{int}})\}$.

**Lemma 1.** VERTEX COVER and INDEPENDENT SET are glueable with 2 colors and DOMINATING SET is glueable with 4 colors.

**Proof.** For VERTEX COVER, we use the color sets $C^\ell_i := \{i^\ell\}$ for $\ell \in \{\text{int}, \text{ext}\}$ and $i = 0, 1$. The function $h$ is $h(x, y, z) = x + y + z$. The subgraphs $G^{\text{ver}}(\chi^{\text{int}})$ for $X \in \{A_1, S, A_2\}$ and $\chi^{\text{int}} : S \to C^0_0 \cup C^1_1$ are $G^{\text{ver}}(\chi^{\text{int}}) := G^{\text{ver}}[X]$. In this way, the subroutine $\text{Eval}_S(\chi^{\text{int}})$ checks whether the coloring $\chi^{\text{int}}$ yields a vertex cover on $G^{\text{ver}}[S]$ and the subroutines $\text{Eval}_{A_i}(\chi^{\text{int}})$ compute the minimum size vertex cover on $G^{\text{ver}}[A_i]$. However, we still need to make sure that all edges going from $A_i$ to $S$ are covered. If a vertex in $S$ is assigned a 1\text{int} by $\chi^{\text{int}}$, the incident edges are already covered. In the case of a 0\text{int}-assignment for a vertex $v \in S$, we can color all neighbors in $N(v) \cap A_i$ to belong to the vertex cover. This is done by the following recolorings "$\oplus A_i."$ Define

$$(\chi^{\text{ext}} \oplus A_i, \chi^{\text{int}})(v) = \begin{cases} 0\text{int} & \text{if } \chi^{\text{int}}(v) = 0\text{int}, \\ 1\text{ext} & \text{if } \chi^{\text{int}}(v) = 1\text{int} \text{ or } \exists w \in N(v) \text{ with } \chi^{\text{int}}(w) = 0\text{int}, \\ \#, & \text{otherwise}. \end{cases}$$

By this recoloring definition, an edge between a separator vertex and a vertex in $A_i$, which is not covered by the separator vertex (due to the currently considered internal coloring) will be covered by the vertex in $A_i$. Our above reasoning shows that – with these settings – equation (I) is satisfied.

INDEPENDENT SET is shown to be glueable with 2 colors by a similar idea.

Regarding DOMINATING SET, we refer to the full paper. □

We want to mention in passing that – besides the problems given in Lemma I – many more select&verify problems are glueable. In particular, this is true for the weighted versions and variations of the above mentioned problems.

## 4 Fixed Parameter Divide and Conquer Algorithms

For the considerations in this sections, let us fix a graph class $G$ for which a $\sqrt{\cdot}$-separator theorem with constants $\alpha$ and $\beta$ is known. Then, we consider a select&verify graph problem $G$ defined by $(P, \text{opt})$ that is glueable with $\sigma$ colors.

### 4.1 Using Glueability for Divide and Conquer

The evaluation of the term $\text{opt}_{x \in \{0, 1\}^n} P_G(x)$ can be done recursively as follows. Start the computation with $\text{opt}\{P_G(x) \mid x \in \{0, 1\}^n\} = \text{opt}\{P_G^{\text{ver}}(x) \mid G^{\text{sel}}) \mid x \in \{0, 1\}^n, x \sim \chi^0_0\}$, where "$\chi^0_0 \equiv \#$" and $G^{\text{ver}} = G^{\text{sel}} = G$.

When $\text{opt}_{x \in \{0, 1\}^n, x \sim \chi^0_0} P_G^{\text{ver}}(x) \mid G^{\text{sel}}$ needs to be calculated for some $G^{\text{sel}}, G^{\text{ver}} \subseteq G$, and an external coloring $\chi^{\text{ext}} : V(G) \to C^0_0 \cup C^1_1 \cup \{\#\}$, we do the following:
1. If $G^{ver}$ has size greater than some constant $c$, then find a $\sqrt{\cdot}$-separator for $G^{ver}$ with $V(G^{ver}) = A_1 + S + A_2$.

2. For all internal colorings $\chi^{int}$ of $S$ with $\chi^{ext} \rightsquigarrow \chi^{int}$ do:
   (a) Determine $\text{Eval}_{A_1}(\chi^{int})$ recursively for $i = 1, 2$.
   (b) Determine $\text{Eval}_{S}(\chi^{int})$.

3. Return $\text{opt}_{\chi^{int}} \cdot \chi^{ext} \rightsquigarrow \chi^{int} \cdot h(\text{Eval}_{A_1}(\chi^{int}), \text{Eval}_{S}(\chi^{int}), \text{Eval}_{A_2}(\chi^{int}))$.

The size of the subproblems, i.e., the size of the graphs $G^{ver}_{A_i}(\chi^{int})$ which are used in the recursion, plays a crucial role in the analysis of the running time of this algorithm.

**Definition 2.** A glueable select&verify problem is called slim if the subgraphs $G^{ver}_{A_i}(\chi^{int})$ are only by a constant number of vertices larger than $G^{ver}[A_i]$, i.e., if there exists a $\eta \geq 0$ such that $|V(G^{ver}_{A_i}(\chi^{int}))| \leq |A_i| + \eta$ for all internal colorings $\chi^{int} : S \rightarrow C^{int}$.

Note that the proof of Lemma 1 shows that both VERTEX COVER and INDEPENDENT SET are slim with $\eta = 0$, whereas DOMINATING SET is not, as exhibited in the paper’s long version. The following proposition gives the running time of the above algorithm. In order to assess the time required for the above given divide and conquer algorithm, we use the following abbreviations for the running times of certain subroutines: $T_S(n)$ denotes the time to find a separator in an $n$-vertex graph from class $\mathcal{G}$. $T_M(n)$ denotes the time to construct the modified graphs $G^{ver}_{X}(\chi^{int}) \in \mathcal{G}$ and the modified colorings $(\chi^{ext} \oplus_X \chi^{int})(\text{for } X = \{A_1, S, A_2\}$ and each internal coloring $\chi^{int}$ with $\chi^{ext} \rightsquigarrow \chi^{int}$) from an $n$-vertex graph from class $\mathcal{G}$. $T_E(m)$ is the time to evaluate $\text{Eval}_{S}(\chi^{int})$ for all $\chi^{int}$, $\chi^{ext} \rightsquigarrow \chi^{int}$, in a separator of size $m = \beta \sqrt{n}$. $T_g(n)$ is the time for gluing the results obtained by two sub-problems each of size $O(n)$. In the following, we assume that all these functions are polynomials.

**Proposition 1.** For every $G \in \mathcal{G}$, $\text{opt}_{x \in \{0,1\}^n} P_G(x)$ can be computed in time

$$c(\alpha', \beta, \sigma) \sqrt{n} q(n), \text{ where } c(\alpha', \beta, \sigma) = \sigma^{\beta/(1-\sqrt{\alpha'})}.$$  

Here, $\alpha' = \alpha + \epsilon$ for any $\epsilon \in (0, 1-\alpha)$ and the running time analysis only holds for $n \geq n_0(\epsilon)$, and $q$ is some polynomial. If, however, $\mathcal{G}$ is slim or the $\sqrt{\cdot}$-separator theorem yields cycle separators, then the running time for the computation is $c(\alpha, \beta, \sigma) \sqrt{n} q(n)$, which then holds for all $n$.

**Proof.** (Sketch) Let $T(n)$ denote the time to compute $\text{opt}_{x \in \{0,1\}^n} x \sim \chi^{ext} P_G^{ver}(x \mid G^{sel})$ for a graph $G^{ver} = (V^{ver}, E^{ver})$ with $n = |V^{ver}|$ (where $\chi^{ext} : V(G) \rightarrow C_0^{ext} + C_1^{ext} + \{\#\}$ is some external coloring and $G^{sel}, G^{ver} \subseteq \mathcal{G}$). In the case of the existence of a cycle separator theorem or if the problem is slim, the recurrence we have to solve in order to compute an upper bound on $T(n)$ then reads as follows:

$$T(n) \leq \sigma^{\beta \sqrt{n}} \cdot 2 T(\alpha n) \cdot (T_M(n) + T_E(\beta \sqrt{n}) + T_g(\alpha n + \beta \sqrt{n})) + T_S(n).$$

$$= T_{M,E,g}(n)$$
The functions $T_{M,E,g}(n)$ and $T_S(n)$ are polynomials. The solution is given by $T(n) \leq \sigma^{(\beta/(1-\sqrt{\alpha}))} \sqrt{n} q(n)$ for some polynomial $q(\cdot)$. In the general case, from the definition of glueability, we have that the size of the two remaining subproblems to be solved recursively is $|\mathcal{V}(G^{\text{ver}}_{A_1}(\chi^{\text{int}}))| \leq \alpha n + \beta \sqrt{n}$ for each $\chi^{\text{int}}$ that preserves $\chi^{\text{ext}}$. Since with $\alpha' = \alpha + \epsilon$, for some $\epsilon \in (0, 1 - \alpha)$, the inequality $\alpha n + \beta \sqrt{n} \leq \alpha'n$ holds for sufficiently large $n$, the result follows.

4.2 How (Linear) Problem Kernels Help

Proposition together with the existence of problem kernels yields:

**Theorem 1.** Suppose that $\mathcal{G}$ admits a problem kernel of polynomial size $p(k)$ on $\mathcal{G}$ computable in time $T_K(n, k)$. Then, there is an algorithm to decide $(G, k) \in \mathcal{G}$, for a graph $G \in \mathcal{G}$, in time

$$c(\alpha', \beta, \sigma) \sqrt{p(k)} q(k) + T_K(n, k), \quad \text{where} \quad c(\alpha', \beta, \sigma) = \sigma^{\beta/(1-\sqrt{\alpha})},$$

(2)

and $\alpha' = \alpha + \epsilon$ for any $\epsilon \in (0, 1 - \alpha)$, holding only for $n \geq n_0(\epsilon)$, where $q(\cdot)$ is some polynomial. If, however, $\mathcal{G}$ is slim or the $\sqrt{\cdot}$-separator theorem yields cycle separators, then the running time for the computation is $c(\alpha, \beta, \sigma) \sqrt{p(k)} q(k) + T_K(n, k)$, which then holds for all $k$.

In particular, this means that for glueable select&verify problems for planar graphs that admit a linear problem kernel of size $dk$, we get an algorithm of running time $O(c(\alpha', \beta, \sigma,d) \sqrt{k} q(k) + T_K(n, k))$, where $c(\alpha', \beta, \sigma,d) = \sigma^{\beta/(1-\sqrt{\alpha})}$.

Since vertex cover is a slim problem, Theorem yields a $c\sqrt{gk}$-algorithm for $\mathcal{G}_g$, where $\mathcal{G}_g$ denotes the class of graphs of genus bounded by $g$, see [5].

4.3 Towards Avoiding (linear) Problem Kernels: The Core Concept

We are going to introduce the novel notion of problem cores, which is closely related to that of problem kernels, but seemingly “incomparable” and tailored towards unweighted minimization select&verify problems. The idea is to restrict (only) the size of the “selection space”, while – unlike in the setting of problem kernels – the whole (possibly large) problem instance may be still used for “checking”.

**Definition 3.** Consider an unweighted select&verify minimization graph problem $\mathcal{G}$ specified by $(P, \min)$. A corer of size $p(k)$ is a polynomial time computable mapping $\phi : ((V, E), k) \mapsto V_c$ satisfying $|V_c| \leq p(k)$, and $\exists \beta(x) = (x_1, \ldots, x_{|V|}) \in \{0, 1\}^{|V|}(P_G(\beta{x}) \leq k \wedge \{v_i \in V \mid x_i = 1\} \subseteq V_c)$. The set $V_c$ is also called the problem core of $\mathcal{G}$. If $p(k) = ak$, we call $\phi$ a linear corer. In this case $V_c$ is called a factor-a problem core.

Note that weighted minimization problems could also be treated similarly at the expense of further technical complications. Having a problem core automatically makes a select&verify problem a “simple” one: For the problem core $V_c$, which can be computed in polynomial time, it is enough to check all $k$-element subsets, giving fixed parameter tractability. Stirling’s formula yields:
Lemma 2. If a select&verify problem $G$ has a size $ak$ problem kernel or if the core is a factor-a core, then there is a $\min\{(ea)^k, 2^ak\}$-algorithm” for $G$.  

Even though there seems to be no general interrelation between problem kernels and cores, for our purposes, the different concepts can be interchanged:

Theorem 2. Let $G$ be an $n$-vertex-graph from a graph class $\mathbb{G}$ for which a $\sqrt{k}$-separator theorem for weighted graphs is known with constants $\alpha, \beta$. Suppose that $G$ admits a corer of size $p(k)$, which can be computed in polynomial time $T_C(n)$. Then, there is an algorithm to decide $(G, k) \in G$, for $G \in \mathbb{G}$, in time

$$c(\alpha', \beta, \sigma)\sqrt{p(k)}q(k) + T_C(n),$$

and $\alpha' = \alpha + \epsilon$ for any $\epsilon \in (0, 1 - \alpha)$, holding only for $n \geq n_0(\epsilon)$. If, however, $G$ is slim or the $\sqrt{k}$-separator theorem yields cycle separators, then the time for the computation is $c(\alpha, \beta, \sigma)\sqrt{p(k)}q(k) + T_C(n)$, which holds for all $k$.

Proof. (Sketch) Consider $G = (V, E) \in \mathbb{G}$. The algorithm proceeds as follows:

1. Compute a core $V_c \subseteq V$ containing at most $p(k)$ vertices in time $T_C(n)$.
2. Then, $G[V_c] \in \mathbb{G}$ is the graph from which vertices have to be selected.
3. Find an optimal $x$ satisfying $P_G(x | G[V_c])$ by applying the algorithm outlined before Prop. One modification of this algorithm, however, is necessary: we do not use a separator theorem in step of that algorithm for a separation of $G^\text{ver}$, but we add weights to $G^\text{ver}$ as follows: $V_c$ induces a weight function $\omega$ on $G^\text{ver}$ by letting $\omega(v) = 1/|V_c|$, if $v \in V_c$ and $\omega(v) = 0$, otherwise. Then, we apply a separator theorem for weighted graphs to $G^\text{ver}$ with the weight function $\omega$. The constants $c(\alpha, \beta, \sigma)$ and $c(\alpha', \beta, \sigma)$ then are derived as in Prop.  

5 Conclusion and Further Results

Summary. We exhibited how to use separator theorems for obtaining $c\sqrt{k}$-fixed parameter divide and conquer algorithms. We defined “glueable select&verify problems,” capturing graph problems such as VERTEX COVER and DOMINATING SET, as a problem class that allows for a divide and conquer approach on certain graph classes. Admittedly, the constants within these algorithms are still rather huge. For example, in the case of PLANAR VERTEX COVER, Theorem yields a $37181\sqrt{k}$-algorithm. In the long version, we elaborate on several ideas to overcome this weakness:

Further results (see the full version of the paper). Firstly, we analyze how Lipton and Tarjan proved their famous planar separator theorem; basically, the proof consists of two steps: in the first step, the given graph is thinned into pieces of “small” radius, and in the second step, a special (cycle) separator lemma for planar graphs with bounded radius is used. The such obtained separator therefore consists of two parts. Since after one application of Lipton and Tarjan’s separator theorem (in a divide and conquer algorithm as described in Section 4),
the remaining graph pieces still have relatively small radius, one could avoid the first thinning step. Iterating this idea in an optimized fashion, one gets, e.g., in the case of PLANAR VERTEX COVER, a $8564\sqrt{k}$-algorithm. It is a challenge if better algorithm bounds are obtainable by using other separator theorems.

Secondly, we discuss the idea of stopping the recursion before having graph parts of constant size and then applying, e.g., elaborated search tree algorithms to these small parts. More precisely, a divide and conquer algorithm would once use Lipton and Tarjan’s planar separator theorem for slicing the graph into pieces of small radius and then use only the mentioned special (cycle) separator lemma in the remainder of the recursion, until all graph pieces are sufficiently small. In this way, one gets $c^{k^{2/3}}$-algorithms with reasonable small constants $c$. For example, we derive a $7.7670^{k^{2/3}}$-algorithm for PLANAR VERTEX COVER.

**Future research.** We briefly sketch three lines of future research. (1) An alternative idea in order to lower the involved constants would be to devise new separator theorems with constants $\alpha$ and $\beta$, not only concentrating on bringing down $\beta$ for fixed $\alpha$ (as, e.g., done for $\alpha = 2/3$ in [6]), but on minimizing the function $\beta/(1 - \sqrt{\alpha})$. (2) It is an issue of future research to further investigate the newly introduced concept of cores. For example: is there a linear size core (or kernel) for PLANAR DOMINATING SET? (3) Finally, this paper (together with [12]) might be a stimulus to investigate the structure “within” FPT in order to distinguish problems allowing for $c^{\sqrt{k}}$-algorithms from problems which seem to allow only for $c^k$-algorithms. Is it possible to develop a reasonable fine-grained structural theory of FPT? Note that Cai and Juedes [3] very recently showed that for a list of parameterized problems (e.g., for VERTEX COVER on general graphs) $c^{o(k)}$-algorithms are impossible unless FPT = W[1].

**References**

On Assigning Prefix Free Codes to the Vertices of a Graph

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Abstract. For a graph G on n vertices, with positive integer weights w1,...,wn assigned to the n vertices such that, for every clique K of G, \( \sum_{i \in K} \frac{1}{2w_i} \leq 1 \), the problem we are interested in is to assign binary codes C1,...,Cn to the vertices such that Ci has wi (or a function of wi) bits in it and, for every edge \( \{i, j\} \), Ci and Cj are not prefixes of each other. We call this the Graph Prefix Free Code Assignment Problem. We relate this new problem to the problem of designing adversaries for comparison based sorting algorithms. We show that the decision version of this problem is as hard as graph colouring and then present results on the existence of these codes for perfect graphs and its subclasses.

1 Introduction

Our study of the Graph Prefix Code Assignment problem is motivated by the question of designing adversaries for sorting under partial information (discussed below). This motivation was suggested by Subrahmanyam and Dubhashi [1].

The problem of sorting under partial information is: Given a partial order P on a set X = \{x1,...,xn\}, determine a hidden linear extension of P using questions of the form is \( x_i < x_j \)? An algorithm to solve this problem is said to sort P by comparisons. It is also referred to as a comparison based algorithm for sorting. In this paper we only consider comparison based sorting algorithms.

A comparison based sorting algorithm can be seen as a game between two players D and H. In this game D has a partial order P and H holds a linear extension of P, T, hidden from D. D tries to discover T by asking questions about the relation in T (comparisons) between pairs of elements. D has to make \( \log e(P) \) comparisons to discover T where, \( e(P) \) is the number of linear extensions of P. An interesting question here is the issue of designing an adversary who answers on-line, the comparison queries of an algorithm, so as to force the algorithm to use \( \log e(P) \) comparisons. This problem was solved by Brodal et

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al. in [7] for the case when $P$ is the anti-chain on $X$. They design a very simple and elegant adversary that answers the comparison queries of any sorting algorithm to force it to use $\Omega(\log n!)$ comparisons. For the general case when there is no restriction on $P$, such an adversary was designed by Kim and Kahn in [2] using a relation they prove between $e(P)$ and the entropy of the comparability graph of $P$(defined below). The adversary answers each comparison query depending on the outcome of an entropy calculation (depending on $P$ and answers to past queries). A question that remains open in this area, posed in [1], is to design an adversary for sorting under partial information that avoids the entropy calculations that are an integral part of the adversary designed in [2].

In this paper we present the relation between the design of adversaries for sorting under partial information and the Graph Prefix free Codes assignment problem. We present the first results on the prefix free codes problem. Our results are stated after the preliminaries in section 2.

2 Preliminaries and Our Results

A collection of positive integers $w_1, \ldots, w_r$ is said to satisfy Kraft’s Inequality if $\sum_{i=1}^{r} \frac{1}{2^{w_i}} \leq 1$.

Throughout the paper, $G$ is an undirected graph with positive integer weights $w_1, \ldots, w_n$ assigned to its vertices $v_1, \ldots, v_n$. The weights satisfy the property that for every clique in $G$, the weights assigned to its vertices satisfy Kraft’s inequality. The weights assigned to the vertices of a graph $G$ are said to satisfy Kraft’s inequality if the induced weight assignment to every clique in $G$ satisfies Kraft’s inequality. Unless mentioned otherwise the weight assignment $\langle w_1, \ldots, w_n \rangle$ to vertices $\langle v_1, \ldots, v_n \rangle$ implies $v_i$ is assigned the weight $w_i$. A similar convention holds for the code assignment $\langle C_1, \ldots, C_n \rangle$. Two codes are prefix free if neither is a prefix of the other.

Prefix Free Code Assignment Problem for Graphs: For a given positive integer $c$, a graph $G$ with weights $w_1, \ldots, w_n$ assigned to vertices $v_1, \ldots, v_n$ satisfying Kraft’s inequality, does there exist an assignment of binary codes $C_1, \ldots, C_n$ to the vertices $v_1, \ldots, v_n$ such that number of bits in $C_i$ is less than or equal to $w_i + c$, and for every edge $\{i, j\}$ in $G$, and such that $C_i$ and $C_j$ are not prefixes of each other?

We call such a code a $+c$ prefix free code for $\langle G, \langle w_1, \ldots, w_n \rangle \rangle$. We use $\langle G, \langle w_1, \ldots, w_n \rangle \rangle$ to denote $G$ with the weights $\langle w_1, \ldots, w_n \rangle$ assigned to its vertices.

Prefix Free Codes and Tree Embeddings: Here we present the relation between prefix free codes and a partial order. Let $T$ denote the infinite binary tree. With each node of $T$, associate a binary string in a natural way: the root is associated with the empty string, and recursively, if a node is associated with a string $w$, its children are associated with the sequences $w0$ and $w1$. $T$ can
be seen as a partial order with the elements related by the ancestor-descendant relationship. Consequently, it follows that two codes are prefix free if and only if they form an anti-chain in $T$. Therefore, our question boils down to assigning vertex $v_i, i \in \{1, \ldots, n\}$, to a node in the tree $T$ at level $w_i + c$ or lesser such that the nodes assigned to adjacent vertices in the graph form an anti-chain in $T$. This problem is now a natural variant of the Graph Vertex Colouring problem. In the Graph Colouring problem, the colours or codes to be assigned to the vertices come from the partial order in which any two distinct elements are unrelated. Therefore, two adjacent vertices must be given colours that are unrelated in this partial order. On the other hand, in the Prefix Free code assignment problem, we are given a tree partial order, and weights to the vertices. The code assignment criterion is that two adjacent vertices get unrelated codes in the partial order and, there is a restriction on the layers where the codes come from.

Definitions Related to Partial Orders and Graph Entropy: The comparability graph, $G(P)$, of a partial order $P$ on a set $X = \{x_1, \ldots, x_n\}$, is an undirected graph with $X$ as its vertex set, and $\{x_i, x_j\} \in E(G(P))$ iff $x_i$ and $x_j$ are comparable in $P$.

We write $P(x < y)$ to denote the partial order obtained from $P$ by adding the relation $x < y$. We refer only to a partial order without referring to the set on which it is defined when there is no confusion about the underlying set.

The vertex packing polytope, $VP(G)$, of a graph $G$ with $V$ as a vertex set, is the convex hull of the set of indicator vectors of stable (independent) sets of $V$. Recall that the indicator vector, $I_A$, of a set $A \subseteq V$, is a $|V|$-dimensional vector in which $(I_A)_i = 1$ if the $i$-th vertex of $V$ is in $A$. Otherwise, it is zero.

The graph entropy of $G$ with respect to a probability distribution $p$ on $V(G)$ is

$$H(G, p) = \min_{a \in VP(G)} - \sum_{i=1}^{n=|V|} p_i \log a_i$$

In this paper we are interested only in the case when $p$ is the uniform distribution on the set of vertices and therefore, entropy is denoted by $H(G)$. The entropy of the comparability graph of a partial order $P$ is denoted by $H(P)$. We will use the phrase entropy of a partial order $P$ to mean the entropy of the comparability graph of $P$.

2.1 Our Results

We first present the relation between the prefix free codes problem and the problem of designing adversaries for sorting under partial information. Using the similarity of the prefix free code assignment problem to graph colouring, we show that

- For an input $\langle G, \langle w_1, \ldots, w_n \rangle \rangle$ where $w_1, \ldots, w_n$ satisfy the Kraft’s inequality, there is no polynomial time algorithm to decide if there is a +0-prefix free code unless $P = NP$. 
We address the problem of the existence of these codes and prove that

- For every positive integer \(c\), there is a graph \(G\), and a vertex weight assignment \(\langle w_1, \ldots, w_n \rangle\) such that there is no \(+c\)-prefix free code.

We then observe that the graphs we construct in the above proof are imperfect graphs, that is, graphs for which there is an induced subgraph whose clique number and chromatic number are different. So we concentrate our exploration on perfect graphs. This direction is also consistent with our goal of designing adversaries for sorting where, the graphs that arise are comparability graphs (this class of graphs are known to be perfect).

- We first construct a perfect graph along with vertex weights for which there is no \(+0\)-prefix free code. We then show that for any perfect graph with weights \(\langle w_1, \ldots, w_n \rangle\) satisfying Kraft’s inequality, there is a prefix free code assignment whose code lengths are within a multiplicative factor of 2 from the specified weights.

We then investigate the problem for specific classes of perfect graphs namely chordal graphs. Specifically,

- we construct a chordal graph \(G\) and associated vertex weights \(w_1, \ldots, w_n\) such that \(\langle G, \langle w_1, \ldots, w_n \rangle \rangle\) does not have a \(+1\)-prefix free code.

**Plan of the Paper.** In Section 3 we present the relation between designing adversaries for sorting under partial information and the prefix free code assignment problem. In Section 4 we present the hardness result for the prefix free code assignment problem. In Section 4.1 we prove non existence results for \(+c\)-prefix free codes for general graphs. In Section 4.2 we study the problem in relation to the perfect graphs. Section 5 presents our results on chordal graphs.

### 3 Motivation for the Prefix Free Code Assignment Problem

In this section we present the connection to the problem of designing adversaries for sorting under partial information. We first present the adversaries due to Brodal [7] and Kim and Kahn [2]. We then point out a method for designing an adversary, which leads to the prefix free code assignment problem. To make the distinction between the two adversaries, we refer to the adversary of Brodal as the adversary for sorting and the adversary of Kim and Kahn as the adversary for sorting under partial information.

#### 3.1 Adversary for Sorting

The adversary that we present is one that forces any sorting algorithm to use \(\Omega(\log n!)\) comparisons.
The Adversary. We describe the adversary of [7] for answering comparisons made by a sorting algorithm. At the outset (before any comparisons are made), the adversary places all the elements of $X$ at the root of an infinite binary tree. For $x \in X$, let $v(x)$ denote the node of the tree to which $x$ has been assigned. The adversary answers a comparison between elements $x, y \in X$ based on the follows cases:

- $v(x) = v(y)$: Then $x$ is moved to the left child of $v(x)$ and $y$ to the right child and the adversary answers $x < y$.
- $v(x)$ is a descendant of $v(y)$: $y$ is moved to the unique child of $v(y)$ that is not an ancestor of $v(x)$. If this child is the left child then the adversary answers $y < x$ and if it is a right child it answers $x < y$.
- $v(x)$ and $v(y)$ are unrelated: If $v(x)$ is visited before $v(y)$ in a pre-order traversal of the tree, the adversary answers $x < y$ and otherwise the adversary answers $y < x$.

It is proved in [7] that this adversary forces any sorting algorithm to use at least $\frac{n \log n}{2}$ comparisons. The key ingredients in the proof of this claim are the following:

- In the worst case, each comparison pushes two elements down by one level in the tree.
- At the end of the algorithm, each node of the tree will contain at most one element.
- Sum of the distances from the root to nodes that contain an element of $X$ is at least $n \log n$ and this is equal to sum of the number of levels each element has been pushed down. This statement is the most important ingredient in the proof of the claim.
- Each comparison contributes at most two to the sum in the previous item. Hence, the number of comparisons made is at least $\frac{n \log n}{2}$.

3.2 Adversary for Sorting under Partial Information

Here we describe the adversary for algorithms that discover a linear extension of a partial order $P$ that is given to them. The adversary described below answers the comparison queries on-line, in a manner so as to force the algorithm to make at least $\log e(P)$ comparisons. The adversary is based on the following two theorems proved in [2] which relate the number of linear extensions of $P$ and $H(P)$.

**Theorem 1.** For any partial order $P$,

$$n(\log n - H(P)) \geq \log e(P),$$

$$\log e(P) \geq \max\{\log n! - nH(P), Cn(\log n - H(P))\},$$

where $C = \frac{1}{1+7\log e}$.

**Theorem 2.** For any partial order $P$ and any incomparable elements $x, y \in P$,

$$\min\{H(P(x < y)), H(P(y < x))\} \leq H(P) + \frac{2}{n}.$$
The Adversary. We describe the answers the adversary gives to a comparison query made by an algorithm that knows a partial order \( P \). The adversary answers a comparison between elements \( x, y \) based on the following cases:

- \( x \) and \( y \) are comparable in \( P \): The adversary answers exactly the relation between \( x \) and \( y \) in \( P \).
- \( x \) and \( y \) are incomparable in \( P \): The adversary applies theorem 2. The adversary answers \( x < y \) if \( H(P(x < y)) = \min\{H(P(x < y)), H(P(y < x))\} \). Otherwise the adversary answers \( y < x \).

In the first case, after getting the answer from the adversary, the algorithm still knows only \( P \). Whereas, in the second case, the algorithm knows a new partial order which is either \( P(x < y) \) or \( P(y < x) \) depending on the answer given by the adversary. It is proved in [2] that this adversary forces any algorithm for sorting under partial information to use \( \Omega(\log e(P)) \) comparisons. The key ingredients in the proof of this claim are the following:

- The entropy of the comparability graph of a total order is \( \log n \). This is because the comparability graph of a total order is a clique.
- By theorem 2, each comparison yields a new partial order whose entropy is at most \( \frac{2}{n} \) more than the entropy of the current partial order.
- The above two observations yield a lower bound on the number of comparisons, namely, \( (\log n - H(P)) \frac{n}{2} \). From theorem 1 this is at least \( \frac{\log e(P)}{2} \).

3.3 A Simpler Adversary for Sorting Partial Orders and the Prefix Free Code Assignment Problem

We now describe a new adversary for sorting under partial information. This adversary starts by placing the elements of \( X \) at nodes in the infinite binary tree. The placement is based on the partial order \( P \). The answers to the comparison queries are given using the rules of the adversary for sorting. We state the following theorem that describes the placement of the elements of \( X \) in the tree:

**Theorem 3.** Let \( P \) be a partial order on a set \( X, |X| = n \). Let \((a_1, \ldots, a_n)\) be a point in \( VP(G(P)) \) such that \( H(G(P)) = \sum_{i=1}^{n} -\log a_i \). Let every element of \( X \) be assigned to a corresponding node of the infinite binary tree satisfying the following properties (tree embedding condition):

- For every \( x, y \in X \), if \( x \) and \( y \) are comparable in \( P \), then the nodes in the tree assigned to \( x \) and \( y \) are unrelated in the tree partial order.
- Element \( i \in X \) is assigned to a node in the infinite binary tree at distance \( \lceil -\log a_i \rceil \) from the root.

Consider an adversary who:

- Starts with the above placement of elements in the infinite binary tree.
- Answers the comparison queries of any algorithm sorting \( P \) using the rules of the adversary for sorting.
This adversary will force any algorithm for sorting under partial information to use at least $\frac{\log e(P) - n}{2}$ comparisons.

Proof. As in the analysis of the adversary for sorting, when the sorting algorithm ends, the adversary would have placed each element of $X$ to a tree node bijectively. The sum of distances from the root to nodes containing elements of $X$ is $n \log n$. The distance from root of an element finally consists of two components: its distance from the root in the tree configuration the adversary starts with, and the distance increase due to comparisons. For each comparison, the contribution to the sum of distances increases by at most 2. Therefore, it follows that $n \log n \leq \sum_{i=1}^{n} \lceil -\log a_i \rceil + 2 \times \text{number of comparisons}$. Using the fact that $H(P) = \frac{1}{n} \sum_{i=1}^{n} -\log a_i$, simple arithmetic and theorem 1 it follows that number of comparisons is at least $\frac{\log e(P) - n}{2}$.

This lower bound on the number of comparisons is of significance for partial orders for which $e(P)$ is large.

Remarks on the Tree Embedding Condition in the Theorem: The condition in the theorem requiring that two related elements in $P$ be assigned two unrelated nodes in the tree partial order is same as asking for the codes assigned to them to be prefix free. The question that is of interest is, Does the tree embedding condition hold for every partial order $P$ with respect to the point in $VP(G(P))$ that achieves $H(P)$? Furthermore, there is no specific reason as to why the tree embedding condition should hold only with respect to a point at which entropy of $P$ is achieved.

Also, from [3] it is true that $VP(G(P))$ consists of exactly those vectors $(v_1, \ldots, v_n) \in [0,1]^n$ for which $\sum_{i \in I} v_i \leq 1$ whenever $\{x_i \in X | i \in I \}$ is a clique in $G(P)$. This statement is true for all perfect graphs and $G(P)$ is a perfect graph. The above remarks lead to the following problem:

Let $P$ be a partial order on a set of $n$ elements. Let $w_1, \ldots, w_n$ be positive integer weights assigned to the vertices of $G(P)$ such that $(\frac{1}{2^{w_1}}, \ldots, \frac{1}{2^{w_n}}) \in VP(G(P))$. Do there exist binary codes $C_1, \ldots, C_n$ to the vertices such that $C_i$ has $w_i$ bits in it and, for every edge $\{i,j\} \in G(P)$, $C_i$ and $C_j$ are not prefixes of each other?

The above problem can be asked for any undirected graph. For an undirected graph there is a simple necessary condition for the existence of prefix free codes: the weights $(w_1, \ldots, w_n)$ should satisfy the Kraft’s inequality. If the weights satisfy Kraft’s inequality then $(\frac{1}{2^{w_1}}, \ldots, \frac{1}{2^{w_n}}) \in VP(G(P))$. Therefore, for arbitrary graphs we get the Graph Prefix Free Code assignment problem:

Let $G$ be a graph on $n$ vertices, with positive integer weights $w_1, \ldots, w_n$ assigned to the $n$ vertices such that the Kraft’s inequality is satisfied. Is there an assignment of binary codes $C_1, \ldots, C_n$ to the vertices such that $C_i$ has $w_i$ bits in it and, for every edge $\{i,j\}$, $C_i$ and $C_j$ are not prefixes of each other?
4 Hardness and Existence Results

In this section, we show that there is no polynomial time algorithm to decide if \( \langle G, \langle w_1, \ldots, w_n \rangle \rangle \) has a +0-prefix free code assignment unless \( P = NP \).

**Theorem 4.** Given a graph \( G \), and a weight assignment \( \langle w_1, \ldots, w_n \rangle \) that satisfies the Kraft’s inequality, there is no polynomial (in the size of the graph) time algorithm to decide if \( \langle G, \langle w_1, \ldots, w_n \rangle \rangle \) has a +0-prefix free code, unless \( P = NP \).

**Proof.** The proof is by showing that a polynomial time algorithm to decide the existence of a +0-prefix free code can be used to obtain in polynomial time an \( NP \)-complete problem. The problem we consider is that of finding in polynomial time a constant factor, less than 2, approximation to the minimum colouring of an undirected graph. Our approximation algorithm would find the smallest \( i, 1 \leq i \leq \log n \) such that \( \langle G, \langle i, \ldots, i \rangle \rangle \) has a +0-prefix free code. This prefix free code assignment would be a \( 2^i \) colouring of \( G \). Since \( i \) is the smallest value for which there is a \( 2^i \) colouring, this gives a constant factor, less than 2, polynomial time approximation algorithm.

4.1 Existence Results

For every positive integer \( c \), we observe that there is a graph \( G \) and a vertex weight assignment such that \( \langle G, \langle w_1, \ldots, w_n \rangle \rangle \) does not have a +\( c \) prefix free code.

**Theorem 5.** For every positive integer \( c \), there is a graph \( G \) with vertex weights \( \langle w_1, \ldots, w_n \rangle \) that satisfies Kraft’s inequality, and \( \langle G, \langle w_1, \ldots, w_n \rangle \rangle \) has no +\( c \)-prefix free code.

**Proof.** We observe that there exist graphs in which the largest clique is an edge, but the chromatic number is arbitrarily large. In particular, for any positive integer \( c \), let \( G_c \) be a graph whose largest clique is an edge and the chromatic number (minimum colouring number) is \( 2^{c+2} \). Let the weight assignment of \( G_c \) be \( \langle 1, \ldots, 1 \rangle \). This weight assignment satisfies Kraft’s inequality in \( G_c \). Clearly, \( G \) does not have a +\( c \) prefix free code assignment.

The construction of graphs with a fixed largest clique number and arbitrarily large chromatic number was first given by Lovasz. A construction of such a graph is presented in the book by Bondy and Murthy [4]. Clearly, the graph that we use in the construction is an imperfect graph, that is graphs which have an induced subgraph whose chromatic and clique number are different. We next consider the prefix free code assignment problem for the case when the graphs are perfect (meaning, not imperfect). The exploration on this class of graphs is well motivated given that comparability graphs are perfect graphs.

4.2 Prefix Free Codes for Perfect Graphs

We next show that the perfectness of a graph is insufficient for +0-prefix free codes to exist.
A Perfect Graph with no +0-Prefix Free Code: The graph we present here is an Interval Graph. Interval graphs are intersection graphs of intervals on a line and are known to be perfect [5]. The graph \( I \) that we construct is the graph in Figure 1. It can be shown that \( \langle I, \langle 1, 2, 2, 2, 1 \rangle \rangle \) does not have a +0-prefix free code. This construction shows that there exist perfect graphs and weight assignments that do not have a +0-prefix free code. We improve on this construction and show in the Section 5 a similar result for a subclass of perfect graphs namely chordal graphs. Now we use the fact that the colouring problem can be solved in polynomial time for perfect graphs, to assign a prefix free code whose lengths are within a multiplicative factor of 2.

Prefix Free Codes within a Constant Factor of the Weight for Perfect Graphs Here we show that the satisfaction of Kraft’s inequality by the weights \( \langle w_1, \ldots, w_n \rangle \) for a perfect graph \( G \) is sufficient for a prefix free code \( \langle C_1, \ldots, C_n \rangle \) to exist such that the number of bits in \( C_i \) is at most \( 2w_i \).

Theorem 6. Let \( G \) be a perfect graph with vertex weights \( \langle w_1, \ldots, w_n \rangle \) that satisfy Kraft’s inequality. \( G \) has a prefix free code \( \langle C_1, \ldots, C_n \rangle \) such that the number of bits in \( C_i \) is at most \( 2w_i \). Moreover, such an assignment of prefix free codes can be computed in polynomial time.

Proof. We prove the theorem by constructing a prefix free code. The construction is as follows: Let there be \( r \) distinct weights such that \( w_1 < \ldots < w_r \) among the weights \( \langle w_1, \ldots, w_n \rangle \). Consider the \( r \) induced subgraphs of \( G \), denoted \( G_1, \ldots, G_r \). \( G_i \) consists of vertices whose weight is \( w_i \) in \( G \). By definition, each of the graphs \( G_1, \ldots, G_r \) is perfect. We find an optimal colouring in polynomial time for each \( G_i \). Due to the perfectness of \( G_i \) and the Kraft’s inequality, the optimal colouring of \( G_i \) uses at most \( 2^{w_i} \) colours. This colouring gives a prefix free code for each \( G_i \) in which every code has \( w_i \) bits. Now for each \( i \), we prefix the codes given to the vertices in \( G_i \) by \( i-1 \) 1s followed by a 0. As \( i \leq w_i \), a vertex of weight \( w_i \) gets a code of length at most \( 2w_i \). Clearly, the code assignment is prefix free. The above construction can be run in polynomial time using the polynomial time algorithm for perfect graph colouring [3]. Hence the theorem.
5 Prefix Free Codes for Chordal Graphs

We construct a chordal graph and a weight assignment for which there is no +1-prefix free code. This construction can be extended to give a chordal graph which has no +c-prefix free code for any positive integer c. We present the construction of the graph in a step by step manner.

– Start with a clique on \(2^w\) vertices, with each vertex of weight \(w\). Clearly this weight assignment satisfies Kraft’s inequality. Let \(G_0\) denote this graph.

– For each sub-clique \(K_1\) of \(G_0\) with \(2^w\) vertices, add a clique \(K_2\) with \(2^w - 1\) vertices, each of weight \(\frac{w}{2}\). Each vertex in \(K_1\) is adjacent to every vertex in \(K_2\) that is, \(K_1 \cup K_2\) is a clique. Let \(G_1\) denote the resulting graph. The weight assignment to \(K_1 \cup K_2\) satisfies Kraft’s inequality.

– For each sub-clique \(K_1\) of \(G_1\), that has \(2^w + 1\) vertices such that \(2^w\) of them have weight \(w\) and the rest have weight \(\frac{w}{2}\), add a clique \(K_2\) with \(2^w (1 - 2^w (\frac{1}{2^w} + \frac{1}{2^w}))\) vertices, each of weight \(\frac{w}{4}\). As before, \(K_1 \cup K_2\) forms a clique and the weight assignment satisfies Kraft’s inequality. Let \(G\) denote the resulting graph.

**Theorem 7.** \(G\) is a chordal graph, and given the weight assignment as mentioned above for \(w \geq 8\), there is no +1-prefix free code.

*Proof.* The proof is omitted due to lack of space.

**Summary** We have motivated the graph prefix free coding problem as a natural problem that arises from the issue of designing adversaries for sorting under partial information. The cases of the problem when the graphs are comparability graphs or interval graphs are of specific interest. Positive results in the comparability graph case would yield adversaries for sorting under partial information and positive results in the interval graph case could be applied to the Dynamic Storage Allocation problem \([p226, SR2]\).

**References**

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A New Measure of Edit Distance between Labeled Trees*

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Abstract. One of the most important problem in computational biology is the tree editing problem which is to determine the edit distance between two rooted labeled trees. It has been shown to have significant applications in both RNA secondary structures and evolutionary trees. Another viewpoint of considering this problem is to find an edit mapping with the minimum cost. By restricting the type of mapping, Zhang [7,8] and Richter [5] independently introduced the constrained edit distance and the structure respecting distance, respectively. They are, in fact, the same concept. In this paper, we define a new measure of the edit distance between two rooted labeled trees, called less-constrained edit distance, by relaxing the restriction of constrained edit mapping. Then we study the algorithmic complexities of computing the less-constrained edit distance between two rooted labeled trees. For unordered labeled trees, we show that this problem is NP-complete and even has no absolute approximation algorithm unless P = NP, which also implies that it is impossible to have a PTAS for the problem. For ordered labeled trees, we give a polynomial-time algorithm to solve the problem.

1 Introduction

A labeled tree is a tree whose nodes are labeled from a finite alphabet $\Sigma$. A rooted tree is a tree in which one of the vertices is designated as root. An unordered labeled tree is just a rooted labeled tree. An ordered labeled tree is a rooted labeled tree in which the children of each node are ordered, i.e., if a node has $k$ children, then we can specify them as the first child, the second child, \ldots, and the $k$th child. Unless otherwise stated, all trees we consider in this paper are rooted labeled trees.

Let $\lambda$ be a symbol not in $\Sigma$ and $a, b \in \Sigma \cup \{\lambda\}$. We use $a \rightarrow b$ to denote an edit operation. Given a tree $T$, we consider three kinds of edit operations on $T$.

1. Change (i.e., $a, b \in \Sigma$): means to change a node $a$ to $b$.
2. Deletion (i.e., $a \in \Sigma$, $b = \lambda$): means to delete a node.
3. Insertion (i.e., $b \in \Sigma$, $a = \lambda$): means to add a node.

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\(b = \lambda\): means to delete a node \(a\) such that the children of \(a\) become the children of the parent of \(a\). (3) Insertion (i.e., \(a = \lambda\)): means to insert a node \(b\) as a child of some one node, say \(c\), such that \(b\) becomes the parent of some (consecutive, if \(T\) is ordered) children of \(c\).

Let \(\gamma\) be a cost function which assigns each edit operation \(a \rightarrow b\) a non-negative real number \(\gamma(a \rightarrow b)\) such that (1) \(\gamma(a \rightarrow a) = 0\), (2) \(\gamma(a \rightarrow b) = \gamma(b \rightarrow a)\), and (3) \(\gamma(a \rightarrow c) \leq \gamma(a \rightarrow b) + \gamma(b \rightarrow c)\). The above function \(\gamma\) is called a metric cost function. Let \(S = (s_1, s_2, \ldots, s_k)\) be a sequence of edit operations. \(S\) is called an edit sequence of transforming tree \(A\) to tree \(B\) if there are trees \(T_1, T_2, \ldots, T_k\) such that \(A = T_1, B = T_{k+1}\) and \(T_{i+1}\), where \(1 \leq i \leq k\), is obtained by applying \(s_i\) to \(T_i\). The cost of \(S\) is defined to be \(\gamma(S) = \sum_{i=1}^{k} \gamma(s_i)\).

For any two trees \(T_1\) and \(T_2\), the edit distance between them, denoted by \(D_e(T_1, T_2)\), is defined to be the minimum cost of an edit sequence transforming \(T_1\) to \(T_2\). That is, \(D_e(T_1, T_2) = \min_{S} \{\gamma(S) | S\ \text{is an edit sequence transforming} \ T_1 \ \text{to} \ T_2\}\). The problem of computing \(D_e(T_1, T_2)\) can be considered from a graphical viewpoint, called edit mapping. Let \(V(T_i)\) be the set of nodes of tree \(T_i\). An edit mapping from \(T_1\) to \(T_2\), denoted by \(M_e(T_1, T_2)\), is defined to be a subset of \(V(T_1) \times V(T_2)\) such that for any two \((u_1, v_1)\) and \((u_2, v_2)\) in \(M_e(T_1, T_2)\), the following conditions are satisfied:

1. \(u_1 = u_2\) if and only if \(v_1 = v_2\) (one-to-one mapping).
2. \(u_1\) is an ancestor of \(u_2\) if and only if \(v_1\) is an ancestor of \(v_2\) (i.e., the ancestor-descendant relationship is preserved).
3. (Only for ordered labeled trees) \(u_1\) is to the left of \(u_2\) if and only if \(v_1\) is to the left of \(v_2\) (i.e., the sibling order is preserved).

See Figure 1 for an example of an edit mapping between two ordered labeled trees \(T_1\) and \(T_2\). We simply use \(M_e\) instead of \(M_e(T_1, T_2)\) if there is no confusion. We call a pair \((u, v) \in M_e\) a mapped line and say that \(u\) and \(v\) are covered by \((u, v)\) or \(M_e\). In fact, the edit mapping shows a way to transform \(T_1\) to \(T_2\) as follows. Each mapped line \((u, v)\) indicates that \(u\) should be changed to \(v\) if \(u \neq v\); otherwise, \(u\) remains unchanged. The uncovered nodes of \(T_1\) are to be deleted and the uncovered nodes of \(T_2\) are to be inserted. For example, the edit mapping of Figure 1 corresponds to an edit sequence \((c \rightarrow \lambda, f \rightarrow g, \lambda \rightarrow c)\) to transform \(T_1\) to \(T_2\).

Let \(U\) and \(V\) be the sets of the uncovered nodes in \(T_1\) and \(T_2\), respectively. Then the cost of \(M_e\) is defined to be \(\gamma(M_e) = \sum_{(u, v) \in M_e} \gamma(u \rightarrow v) + \sum_{u \in U} \gamma(u \rightarrow \lambda) + \sum_{v \in V} \gamma(\lambda \rightarrow v)\). In [10], Zhang and Shasha showed that the edit sequence with the minimum cost corresponds to an edit mapping with the minimum cost. In other words, the edit distance between \(T_1\) and \(T_2\) can be restated as \(D_e(T_1, T_2) = \min \{ \gamma(M_e) | M_e\ \text{is an edit mapping between} \ T_1 \ \text{and} \ T_2\}\).

Let \(lca(u, v)\) represent the least common ancestor of nodes \(u\) and \(v\) in a tree. Motivated by [6], Zhang [7] introduced the constrained edit mapping, denoted by \(M_c\), which is an edit mapping satisfying an additional condition that for any triple \((u_1, v_1), (u_2, v_2)\) and \((u_3, v_3)\) in \(M_c\), \(lca(u_1, u_2)\) is a proper ancestor of \(u_3\) if and only if \(lca(v_1, v_2)\) is a proper ancestor of \(v_3\). In [5], Richter defined the
structure respecting mapping, denoted by $M_s$, which is an edit mapping satisfying an additional condition that for any triple $(u_1, v_1), (u_2, v_2)$ and $(u_3, v_3)$ in $M_s$ such that none of $u_1, u_2$ and $u_3$ is an ancestor of one of the others, $lca(u_1, u_2) = lca(u_1, u_3)$ if and only if $lca(v_1, v_2) = lca(v_1, v_3)$. In fact, both the concepts of constrained edit mapping and structure respecting mapping are equivalent. Each one is a restricted variant of the edit mapping which prohibits the edit mappings as shown in Figures 3 and 4. The mapping of Figure 2 is a constrained edit mapping since $T_1$ and $T_2$ are isomorphic. The mapping of Figure 3 is not a constrained edit mapping since $lca(u_1, u_2) = u_4 \neq lca(u_1, u_3) = u_5$, but $lca(v_1, v_2) = u_4 = lca(v_1, v_3)$, or since $lca(u_1, u_2)$ is not a proper ancestor of $u_3$, but $lca(v_1, v_2)$ is a proper ancestor of $v_3$. The mapping of Figure 4 is not a constrained edit mapping since $lca(u_1, u_3) = u_5 = lca(u_2, u_3)$, but $lca(v_1, v_3) = v_5 \neq lca(v_2, v_3) = v_4$, or since $lca(u_2, u_3)$ is a proper ancestor of $u_1$, but $lca(v_2, v_3)$ is not a proper ancestor of $v_1$. The constrained edit distance (or structure respecting distance) between two trees $T_1$ and $T_2$ is defined to be $D_c(T_1, T_2) = \min \{ \gamma(M_c) | M_c$ is a constrained edit mapping between $T_1$ and $T_2 \}$.

In this paper, we define a new variant of the edit mapping, called less-constrained edit mapping, by relaxing the restriction of constrained edit mapping such that the edit mapping of Figure 3 is allowable. The less-constrained edit mapping $M_l$ is defined to be an edit mapping which satisfies an additional condition that for any triple $(u_1, v_1), (u_2, v_2)$ and $(u_3, v_3)$ in $M_l$ such that none of $u_1, u_2$ and $u_3$ is an ancestor of the others, $lca(u_1, u_2) \leq lca(u_1, u_3) = lca(u_2, u_3)$ if and only if $lca(v_1, v_2) \leq lca(v_1, v_3) = lca(v_2, v_3)$. However, the mapping of Figure 4 is not a less-constrained edit mapping since $lca(u_1, u_2) < lca(u_1, u_3) = lca(u_2, u_3)$ and $lca(v_1, v_2) = lca(v_1, v_3) = lca(v_2, v_3)$. We define the less-constrained edit distance between two trees $T_1$ and $T_2$ as $D_l(T_1, T_2) = \min \{ \gamma(M_l) | M_l$ is a less-constrained edit mapping between $T_1$ and $T_2 \}$.

The editing (respectively, constrained editing and less-constrained editing) problem is the problem of computing $D_c(T_1, T_2)$ (respectively, $D_c(T_1, T_2)$ and $D_l(T_1, T_2)$) between two trees $T_1$ and $T_2$. The editing problem is shown to

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**Fig. 1.** The dashed lines denote an edit mapping.

**Fig. 2.** An edit mapping which is constrained and less-constrained.
be NP-complete \cite{11} and even be MAX SNP-hard \cite{9} for two unordered labeled trees, but is solvable in $O(|T_1| \times |T_2| \times \min\{\text{depth}(T_1), \text{leaves}(T_1)\} \times \min\{\text{depth}(T_2), \text{leaves}(T_2)\})$ time for two ordered labeled trees \cite{10}. The constrained editing problem can be solved in $O(|T_1| \times |T_2| \times (\text{degree}(T_1) + \text{degree}(T_2)) \times \log(\text{degree}(T_1) + \text{degree}(T_2)))$ time for unordered labeled trees \cite{8} and in $O(|T_1| \times |T_2|)$ time for ordered labeled trees \cite{7}. In this paper, we show that the less-constrained editing problem for unordered labeled trees is NP-complete and even has no absolute approximation algorithm, unless P = NP. The negative result of absolute approximation also implies that it is very unlikely to have a PTAS for the problem. In fact, we can show this problem to be MAX SNP-hard. However, this result does not necessarily imply that it has no absolute approximation algorithm. For ordered labeled trees, we give an algorithm of $O(|T_1| \times |T_2| \times \text{degree}^3(T_1) \times \text{degree}^3(T_2) \times (\text{degree}(T_1) + \text{degree}(T_2)))$ time to solve it.

2 NP-Completeness Result

LCEP (Less-Constrained Editing Problem on Unordered Labeled Trees)
Instance: Two unordered labeled tree $T_1$ and $T_2$ and a positive integer $K$.
Question: Is the less-constrained edit distance between $T_1$ and $T_2 \leq K$?

X3C (Exact Cover by 3-Sets Problem)
Instance: A finite set $X$ with $|X| = 3n$ and a collection $S$ of 3-element subsets of $X$ with $|S| = m$.
Question: Does $S$ contain an exact cover for $X$, i.e., a subcollection $S' \subseteq S$ such that every element of $X$ occurs in exactly one member of $S$?

Note that X3C is a NP-complete problem \cite{3}. Let $X = \{x_1, x_2, \ldots, x_{3n}\}$ and $S = \{S_1, S_2, \ldots, S_m\}$ be an instance of X3C. Without loss of generality, we assume that $m > n$. Then we transform $X$ and $S$ into an instance of LCEP, $T_1$ and $T_2$, as shown in Figure 3 and 4 respectively. For convenience, we let $S_i = \{t_{i,1}, t_{i,2}, t_{i,3}\}$ for $1 \leq i \leq m$ and call a node with label $\alpha$ an $\alpha$-node.

The root of $T_1$ is an $R$-node which has $m$ $S$-nodes as the children to correspond to $S_1, S_2, \ldots, S_m$ of $S$. Each $S$-node has three children which are labeled $t_{i,1}, t_{i,2}$ and $t_{i,3}$ respectively if it corresponds to $S_i$. The root of $T_2$ is an $R$-node and has $m + 2n$ children in which $3n$ children are labeled $x_1, x_2, \ldots, x_{3n}$ respec-

Fig. 3. An edit mapping which is less-constrained, but not constrained. Fig. 4. An edit mapping which is neither constrained nor less-constrained.
Suppose that \( k < y \) operation with cost 1. That is, the covered children. Clearly, the construction of \( T_1 \) and \( T_2 \) can be accomplished in polynomial time. For convenience, we assume that the cost of each edit operation is 1, i.e., \( \gamma(u \to v) = 1 \) if \( u \neq v \).

**Lemma 1.** Let \( M \) be any less-constrained edit mapping from \( T_1 \) to \( T_2 \). Suppose that there are \( k \), \( 1 \leq k \leq |T_2| \), uncovered nodes in \( T_2 \). Then \( \gamma(M) \geq 3m - 2n + k \).

**Proof.** Let \( U \) and \( V \) be the sets of the uncovered nodes in \( T_1 \) and \( T_2 \), respectively. Then \( |V| = k \). Since \( |T_1| = 4m + 1 \), \( |T_2| = 4m - n + 1 \) and \( M \) is a one-to-one mapping, \( |M| = |T_2| \setminus V = 4m - n - k + 1 \) and \( |U| = |T_1| - |M| = n + k \). For each uncovered node in \( U \), it corresponds to a delete (insert) operation with cost 1. Hence, \( U \) and \( V \) totally contribute a cost \( n + 2k \) to \( \gamma(M) \), i.e., \( \gamma(M) \geq n + 2k \).

If \( k \geq 3m - 3n \), then \( n + 2k \geq 3m - 2n + k \) and hence \( \gamma(M) \geq 3m - 2n + k \). Suppose that \( k < 3m - 3n \). Then \( T_2 \) has at least \( (3m - 3n - k) \) \( y \)-nodes covered by \( M \). Since \( T_1 \) has no \( y \)-node, each covered \( y \)-node corresponds to a change operation with cost 1. That is, the covered \( y \)-nodes in \( T_2 \) contribute a cost at least \( 3m - 3n - k \) to \( \gamma(M) \). Totally, \( U \), \( V \) and the covered \( y \)-nodes in \( T_2 \) contribute a cost at least \( 3m - 2n + k \) to \( \gamma(M) \). Hence, \( \gamma(M) \geq 3m - 2n + k \).

**Theorem 1.** LCEP is an NP-complete problem.

**Proof.** (Sketch) Given an instance \( X \) and \( S \) of X3C, we create an instance \( T_1 \) and \( T_2 \) of LCEP as described in the beginning of this section. First, suppose that \( S \) has an exact cover \( S' \). Then we find a less-constrained edit mapping \( M \) from \( T_1 \) to \( T_2 \) as follows. (1) \( R \)-node of \( T_1 \) is mapped to \( R \)-node of \( T_2 \). (2) If \( S_i \) is in \( S' \), then we map the children of its corresponding \( S \)-node in \( T_1 \) to \( t_{i,1} \)-node, \( t_{i,2} \)-node and \( t_{i,3} \)-node in \( T_2 \). Note that the \( S \)-node is not mapped. (3) If \( S_i \) is not in \( S' \), then the subtree of \( T_1 \) rooted at the corresponding \( S \)-node of \( S_i \) is mapped to a subtree of \( T_2 \) rooted at an \( S \)-node in a way that root is mapped to root and leaf is mapped to leaf. Clearly, the above mapping \( M \) is a less-constrained edit mapping and \( D_l(T_1, T_2) \leq 3m - 2n \).

Conversely, suppose that there is a less-constrained edit mapping \( M \) from \( T_1 \) to \( T_2 \) with cost \( \leq 3m - 2n \). By Lemma 1 all nodes in \( T_2 \) are covered by \( M \) and hence \( T_1 \) contain \( n \) uncovered nodes. Then the uncovered nodes of \( T_1 \) and...
the covered $y$-nodes of $T_2$ contribute a total cost $3m - 2n$ to $\gamma(M)$, which leads to $\gamma(M) = 3m - 2n$. Hence, $R$-node of $T_2$ is mapped to $R$-node of $T_1$, $S$-node of $T_2$ is mapped to $S$-node of $T_1$, and $x_i$-node of $T_2$ is mapped to $x_i$-node of $T_1$, $1 \leq i \leq 3n$. Let $S'$ and $S''$ denote any two $S$-nodes, which are mapped to each other, in $T_1$ and $T_2$, respectively. Then the children of $S''$ (i.e., 3 $y$-nodes) are all mapped to the children of $S'$ since the ancestor-descendant relationship is preserved by $M$. Hence, the uncovered nodes in $T_1$ are $n$ $S$-nodes and their children are mapped to $x_1$-node, $x_2$-node, $\cdots$, $x_{3n}$-node in $T_2$, which implies that these $n$ uncovered $S$-nodes of $T_1$ correspond to an exact cover of $S$. \qed

3 Negative Result for Absolute Approximation

Recall that Zhang and Jiang [9] showed that the editing problem for two unordered labeled trees is MAX SNP-hard by a reduction from the maximum bounded covering by 3-sets, which is known to be MAX SNP-hard [1]. In fact, their proof can be applied to show that the less-constrained editing problem for two unordered labeled trees is MAX SNP-hard. Note that any MAX SNP-hard problem has no PTAS (Polynomial Time Approximation Scheme), unless $P = \text{NP}$ [1], where a problem has a PTAS if for any fixed $\epsilon > 0$, the problem can be approximated within a factor of $1 + \epsilon$ in polynomial time. Hence, it is very unlikely for the less-constrained editing problem to have a PTAS. However, this negative result does not imply that there is no polynomial-time absolute approximation algorithm for the problem. An approximation algorithm is absolute if there exists a constant $c$ such that for every instance $I$ of the problem, the absolute error $|\text{APP}(I) - \text{OPT}(I)| \leq c$, where $\text{APP}(I)$ and $\text{OPT}(I)$ are the approximate and optimal solutions of $I$, respectively [2]. In this section, we propose a technique of tree composition to show that the less-constrained editing problem for two unordered labeled trees has no polynomial-time absolute approximation algorithm, unless $P = \text{NP}$. For convenience, we assume that the cost of each edit operation is 1, i.e., $\gamma(u \to v) = 1$ if $u \neq v$. For any two forests $F_1$ and $F_2$, we use $D_t(F_1, F_2)$ to denote the less-constrained edit distance between them, which is defined analogously to the case of two trees.

**Lemma 2.** Let $F_1$ and $F_2$ be any two forests of unordered labeled trees. Then we have $||F_1| - |F_2|| \leq D_t(F_1, F_2) \leq |F_1| + |F_2|$.

Given any $k$ unordered labeled trees $T_1, T_2, \cdots, T_k$, the composition of $T_1, T_2, \cdots, T_k$, denoted by $T_1 \circ T_2 \circ \cdots \circ T_k$, is defined to be an unordered labeled tree which is obtained from $T_1, T_2, \cdots, T_k$ by introducing a new root $R$ such that the root of each $T_i$, $1 \leq i \leq k$, is a child of $R$ (see Figure 7). For convenience, we define $T_i^k = T_1 \circ T_2 \circ \cdots \circ T_k$ iff $T_1 = T_2 = \cdots = T_k$.

Let $T_1$ and $T_2$ be any two unordered labeled trees and $k$ be any positive integer. Then we construct two new unordered labeled trees $G_1 = (T_1 \circ H)^k$ and $G_2 = (T_2 \circ H)^k$, where $H$ is a complete binary tree with $(2k+1)(|T_1|+|T_2|)$ $y$-nodes (see Figure 8 for an example of $G_1$). In each $G_i$, $i = 1, 2$, the root has $k$ children and the subtree rooted at each such child is $T_i \circ H$. We denote by $F_i$
the collection of $k$ subtrees $T_1 \circ H$ rooted at the children of the root of $G_i$. It is not hard to verify the following two lemmas.

**Lemma 3.** $D_l(G_1, G_2) \leq kD_l(T_1, T_2)$.

**Lemma 4.** Let $M$ be an arbitrary less-constrained edit mapping with the minimum cost from $G_1$ to $G_2$. Then each subtree of $F_1$ is mapped to exactly one subtree of $F_2$ by $M$.

**Lemma 5.** $D_l(G_1, G_2) \geq kD_l(T_1, T_2)$.

**Proof.** (Sketch) By Lemma 4, each subtree $A = T_1 \circ H$ of $F_1$ is mapped to exactly one subtree $B = T_2 \circ H$ of $F_2$ by an optimal mapping $M$. Moreover, root($A$) and root($B$) are mapped each other. We use $H_A$ and $H_B$ to denote $H$ of $A$ and $H$ of $B$ respectively. Suppose that $H_A$ is not covered by $M$. Then $\gamma(M) \geq |H_A| > kD_l(T_1, T_2)$, a contradiction. Case 1: $H_A$ is mapped to exactly $T_2$. By Lemma 2 $\gamma(M) \geq |H_A| - |T_2| > kD_l(T_1, T_2)$, a contradiction. Case 2: $H_A$ is mapped to exactly $H_B$. The distance between $A$ and $B$ must be at least $D_l(T_1, T_2)$. Case 3: $H_A$ is mapped to both $T_2$ and $H_B$. Then root($H_A$) and $T_1$ are not covered by $M$. Hence, there are at least $|T_1| + 1$ uncovered nodes in $A$ and at least $|H_B| + |T_2| - |H_A \setminus \{\text{root}(H_A)\}| = |T_2| + 1$ uncovered nodes in $B$, i.e., the distance between $A$ and $B$ is at least $|T_1| + |T_2| + 2 > D_l(T_1, T_2)$. □

By Lemmas 3 and 4 we have the following theorem.

**Theorem 2.** $D_l(G_1, G_2) = kD_l(T_1, T_2)$ for any two unordered labeled trees $T_1$ and $T_2$.

**Theorem 3.** If there is a polynomial-time absolute approximation algorithm of determining the less-constrained edit distance between two unordered labeled trees, then $P = NP$.

**Proof.** Suppose that the problem of determining $D_l(T_1, T_2)$ can be approximated by an absolute approximation algorithm $A$ with absolute error $c$ in polynomial time $O((|T_1| + |T_2|)^\beta)$ for some constant $\beta$. Then we let $k = c + 1$ and create $G_1 =$
(T_1 \circ H)^k and G_2 = (T_2 \circ H)^k as described previously with each size O(k^2(|T_1| + |T_2|)). By applying A to G_1 and G_2, we can obtain an approximate solution APP(G_1, G_2) of D_l(G_1, G_2) in time O(k^{2\beta}((|T_1| + |T_2|)^{\beta})), which is polynomial in |T_1| + |T_2| since k and \beta are fixed, such that |APP(G_1, G_2) - D_l(G_1, G_2)| \leq c.

By Theorem 2 we have |APP(G_1, G_2) - kD_l(T_1, T_2)| \leq c \iff |\frac{APP(G_1, G_2)}{k} - D_l(T_1, T_2)| \leq \frac{c}{k}. Let M denote the corresponding mapping of APP(G_1, G_2). Then each subtree T_1 \circ H of F_1 is mapped to exactly one subtree T_2 \circ H of F_2 in M; otherwise, APP(G_1, G_2) \geq |T_1 \circ H| > |H| = (2k + 1)(|T_1| + |T_2|) (where i = 1 or i = 2), which leads to the absolute error |APP(G_1, G_2) - D_l(G_1, G_2)| \geq (k + 1)(|T_1| + |T_2|) > c, a contradiction. We use M' \subseteq M to denote the mapping from T_1 \circ H to T_2 \circ H with the minimum cost among k such mappings and use H_1 and H_2 to denote H's of T_1 \circ H and T_2 \circ H, respectively. Clearly, the cost of M' \leq \frac{APP(G_1, G_2)}{k}. Note that H_1 must be covered by M'. Case 1: H_1 is mapped to exactly T_2. Then APP(G_1, G_2) \geq |H_1| - |T_2| \geq 2k(|T_1| + |T_2|), which leads to |APP(G_1, G_2) - D_l(G_1, G_2)| \geq k(|T_1| + |T_2|) > c, a contradiction. Case 2: H_1 is mapped to exactly H_2. In this case, we use the mapping of M' between T_1 and T_2 as the approximate solution of finding the optimal mapping from T_1 to T_2, i.e., APP(T_1, T_2) \leq \gamma(M') \leq \frac{APP(G_1, G_2)}{k}. Case 3: H_1 is mapped to both T_2 and H_2.

Hence, we find an approximate solution APP(T_1, T_2) of D_l(T_1, T_2) according to APP(G_1, G_2) such that APP(T_1, T_2) \leq \frac{APP(G_1, G_2)}{k}. Then |APP(T_1, T_2) - D_l(T_1, T_2)| \leq |\frac{APP(G_1, G_2)}{k} - D_l(T_1, T_2)| \leq \frac{c}{k} = \frac{c}{c+1} < 1. Since both APP(T_1, T_2) and D_l(T_1, T_2) are integer-valued, APP(T_1, T_2) = D_l(T_1, T_2), which means that the less-constrained editing problem is solvable by A in polynomial time and hence we have P = NP by Theorem 10.

4 A Polynomial-Time Algorithm for Ordered Labeled Trees

We use D(T_1, T_2) instead of D_l(T_1, T_2) if no confusion arises. Given an ordered labeled tree T_i, for convenience, we use the left-to-right postorder numbering to number the nodes of T_i from 1 to |T_i|. Following the notation used in [7,8], we use t_i[k] to denote the kth node of T_i, T_i[k] to denote the subtree rooted at t_i[k] and F_i[k] to denote the ordered forest obtained by deleting t_i[k] from T_i[k]. Note that the left-to-right order of the trees in F_i[k] is significant. For any two forests F_1[i] and F_2[j], we use D(F_1[i], F_2[j]) to denote the less-constrained edit distance between them. Let \emptyset denote the empty tree. Throughout the rest of this section, we let 1 \leq i \leq |T_1| and 1 \leq j \leq |T_2| and let t_1[i_1], t_1[i_2], \ldots, t_1[i_n] be the children of t_1[i] and t_2[j_1], t_2[j_2], \ldots, t_2[j_n] be the children of t_2[j]. Due to the limitation of space, some proofs are omitted in this section.
Lemma 6. (1) \( D(\emptyset, \emptyset) = 0 \), (2) \( D(F_1[i], \emptyset) = \sum_{k=1}^{n_i} D(T_1[i_k], \emptyset) \), \( D(T_1[i], \emptyset) = D(F_1[i], \emptyset) + D(T_1[i], \emptyset) \). (3) \( D(\emptyset, F_2[j]) = \sum_{k=1}^{n_j} D(\emptyset, T_2[j_k]) \), \( D(\emptyset, T_2[j]) = D(\emptyset, F_2[j]) + \gamma(\lambda \rightarrow t_2[j]) \).

Lemma 7. Let \( M \) be a less-constrained edit mapping from \( T_1[i] \) to \( T_2[j] \). (1) If both \( t_1[i] \) and \( t_2[j] \) are not covered by \( M \), then \( M \cup \{(t_1[i], t_2[j])\} \) is a less-constrained edit mapping. (2) If both \( t_1[i] \) and \( t_2[j] \) are covered by \( M \), then \( (t_1[i], t_2[j]) \in M \).

Lemma 8. \( D(T_1[i], T_2[j]) = \min \{D(\emptyset, T_2[j]) + \min_{1 \leq k \leq n_i} \{D(T_1[i], T_2[j_k]) - D(\emptyset, T_2[j_k])\}, D(T_1[i], \emptyset) + \min_{1 \leq k \leq n_i} \{D(T_1[i_k], T_2[j]) - D(T_1[i_k], \emptyset)\}, \gamma(t_1[i] \rightarrow t_2[j]) + D(F_1[i], F_2[j])\} \).

Proof. (Sketch) Let \( M \) be an less-constrained edit mapping with the minimum cost from \( T_1[i] \) to \( T_2[j] \), i.e., \( D(T_1[i], T_2[j]) = \gamma(M) \). By Lemma 7 we can compute \( \gamma(M) \) by just considering the following three cases. Case 1: \( t_1[i] \) is covered by \( M \), but \( t_2[j] \) is not. Case 2: \( t_2[j] \) is covered by \( M \), but \( t_1[i] \) is not. Case 3: \( t_1[i] \) and \( t_2[j] \) are both covered by \( M \).

Suppose that \( M \) is any arbitrary less-constrained edit mapping from \( F_1[i] \) to \( F_2[j] \). In \( F_1[i] \) and \( F_2[j] \), if no tree is mapped to more than one other tree by \( M \), then we call \( M \) a 1-to-1 less-constrained edit mapping; otherwise, we call \( M \) a 1-to-many less-constrained edit mapping. We use \( D_1(F_1[i], F_2[j]) \) and \( D_2(F_1[i], F_2[j]) \) to denote the minimum costs of 1-to-1 and 1-to-many less-constrained edit mappings from \( F_1[i] \) to \( F_2[j] \), respectively. Clearly, we have \( D(F_1[i], F_2[j]) = \min \{D_1(F_1[i], F_2[j]), D_2(F_1[i], F_2[j])\} \).

Lemma 9. \( D_1(F_1[i], F_2[j]) = \min \{D_1(F_1[i], T_2[j_1]) \cdots T_2[j_{n_j-1}] + D(\emptyset, T_2[j_{n_j}]), D_1(T_1[i_1]) \cdots T_1[i_{n_i-1}], F_2[j]) + D(T_1[i_{n_i}], \emptyset), D_1(T_1[i_1]) \cdots T_1[i_{n_i-1}], T_2[j_1] \cdots T_2[j_{n_j-1}] + D(T_1[i_{n_i}], T_2[j_{n_j}])\} \).

Lemma 10. Let \( M \) be a 1-to-many less-constrained edit mapping from \( F_1[i] \) to \( F_2[j] \). If \( T_1[i_k] \) of \( F_1[i] \) \( (T_2[j_k] \) of \( F_2[j]) \) is mapped to more than one tree of \( F_2[j] \) \( (F_1[i]) \), then \( t_1[i_k] \) \( (t_2[j_k]) \) is not covered by \( M \).

Proof. (Sketch) Let \( t_1[a_1] \) and \( t_1[a_2] \) be any two nodes in \( T_1[i_k] \) such that they are mapped to \( t_2[b_1] \) of \( T_2[j_p] \) and \( t_2[b_2] \) of \( T_2[j_q] \) respectively. Suppose that \( t_1[i_k] \) is mapped to \( t_2[b_3] \). Then, due to the ancestor-descendant relationship of \( M \), \( t_2[b_3] \) must be a common ancestor of \( t_2[b_1] \) and \( t_2[b_2] \), which contradicts to the fact that \( T_2[j_p] \) and \( T_2[j_q] \) are two disjoint trees.

Lemma 11. Let \( M \) be a 1-to-many less-constrained edit mapping from \( F_1[i] \) to \( F_2[j] \) and let \( T_1[i_k] \) of \( F_1[i] \) be mapped to \( T_2[j_l] \) \( (T_2[j_k] \) of \( F_2[j]) \), \( 1 \leq l < r \leq n_j \). (1) For any \( 1 \leq p < k \), \( T_1[i_p] \) can only be mapped to \( T_2[j_l] \cdots T_2[j_{l-1}] \). (2) For any \( k < q \leq n_i \), \( T_1[i_q] \) can only be mapped to \( T_2[j_{r+1}] \cdots T_2[j_n] \).
Proof. (Sketch) (1) Let \( t_1[a_1] \) and \( t_1[a_2] \) of \( T_1[i_k] \) be mapped to \( t_2[b_1] \) of \( T_2[j_i] \) and \( t_2[b_2] \) of \( T_2[j_r] \) respectively, and \( a_1 < a_2 \). Let \( t_1[a_3] \) be a node of \( T_1[i_p] \) that is mapped to \( t_2[b_3] \). Then \( a_3 < a_1 < a_2 \) and hence \( b_3 < b_1 < b_2 \). If \( t_2[b_3] \) is a node of \( T_2[j_i] \), then \( \text{lcd}(t_2[b_3], t_2[b_1]) < \text{lcd}(t_2[b_3], t_2[b_2]) = \text{lcd}(t_2[b_1], t_2[b_2]) \) and \( \text{lcd}(t_1[a_3], t_1[a_1]) = \text{lcd}(t_1[a_3], t_1[a_2]) > \text{lcd}(t_1[a_1], t_1[a_2]) \), a contradiction. □

Lemma 12. Let \( M \) be a 1-to-many less-constrained edit mapping from \( F_1[i] \) to \( F_2[j] \) and let \( T_2[j_k] \) of \( F_2[j] \) be mapped to \( T_1[i_l] \cdots T_1[i_r] \), \( 1 \leq l < r \leq n_i \). (1) For any \( 1 \leq p < k \), \( T_2[j_p] \) can only be mapped to \( T_1[i_l] \cdots T_1[i_{l-1}] \). (2) For any \( k < q \leq n_j \), \( T_2[j_q] \) can only be mapped to \( T_1[i_{r+1}] \cdots T_1[i_{n_i}] \).

Lemma 13. \( D_2(F_1[i], F_2[j]) = \)

\[
\min \left\{ \begin{array}{l}
\min_{1 \leq k \leq n_i} \min_{1 \leq i \leq n_i} \left( \sum_{k=1}^{i} D(T_1[i_k] \cdots T_1[i_{k-1}], T_2[j_1] \cdots T_2[j_{i-1}]) + D(T_1[i_k], T_2[j_k]) + \gamma(t_1[i] \to t_2[j]) \right), \right.
\min_{1 \leq k \leq n_j} \min_{1 \leq i \leq n_i} \left( \sum_{k=1}^{j} D(T_1[i_1] \cdots T_1[i_{j-1}], T_2[j_1] \cdots T_2[j_{k-1}]) + D(T_1[i_k], T_2[j_k]) + \gamma(t_1[i] \to t_2[j]) \right). \end{array} \right.
\]

By above lemmas, we design the following algorithm to solve the problem in \( O(|T_1| \times |T_2| \times \text{degree}^3(T_1) \times \text{degree}^3(T_2) \times (\text{degree}(T_1) + \text{degree}(T_2))) \) time.

1: \( D(\emptyset, \emptyset) = 0 \), \( D_1(\emptyset, \emptyset) = 0 \), \( D_2(\emptyset, \emptyset) = 0 \); /* Initialization */

for \( i = 1 \) to \( |T_1| \) do

\( D(F_1[i], \emptyset) = \sum_{k=1}^{i} D(T_1[i_k], \emptyset); \)

\( D(T_1[i], \emptyset) = D(F_1[i], \emptyset) + \gamma(t_1[i] \to \lambda); \)

for \( j = 1 \) to \( |T_2| \) do

\( D(\emptyset, F_2[j]) = \sum_{k=1}^{j} D(\emptyset, T_2[i_k]); \)

\( D(\emptyset, T_2[j]) = D(\emptyset, F_2[j]) + \gamma(\lambda \to t_2[j]); \)

2: for \( i = 1 \) to \( |T_1| \) do /* The computation of \( D(T_1[i], T_2[j]) \) */

for \( j = 1 \) to \( |T_2| \) do

for \( p_1 = 1 \) to \( n_i \) do

for \( p_2 = p_1 \) downto 1 do

for \( q_1 = 1 \) to \( n_j \) do

for \( q_2 = q_1 \) downto 1 do

/* Let \( F_1 = T_1[i_{p_2}] \cdots T_1[i_{p_1}] \) and \( F_2 = T_2[i_{q_2}] \cdots T_2[i_{q_1}] \). */

Compute \( D_1(F_1, F_2) \) by Lemma 8 and \( D_2(F_1, F_2) \) by Lemma 13.

\( D_1(F_1, F_2) = \min\{D_1(F_1, F_2), D_2(F_1, F_2)\}; \)

end for

end for

end for

end for

3: Output \( D(T_1[|T_1|], T_2[|T_2|]) \).
References


A Highly Efficient Algorithm to Determine Bicritical Graphs

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Abstract. In this paper, we show a necessary and sufficient condition which characterizes all bicritical graphs. Using this necessary and sufficient condition we develop a highly efficient algorithm to determine whether a graph is bicritical, of which the time complexity is bounded by $O(|V||E|)$, and this is the best result that has ever been known.

1 Introduction and Terminology

All graphs considered in this paper are finite, undirected, connected and simple. In general, we follow the terminology of [1]. A graph $G$ is said to be bicritical if $G - u - v$ has a perfect matching for each pair of vertices $u$ and $v$ in $G$ such that $u \neq v$. Plummer [10] introduced the concept of $n$-extendable graphs. Let $G$ be a graph on $v$ vertices and $n$ be an integer such that $0 \leq n \leq (v - 2)/2$. Then $G$ is $n$-extendable if every matching of size $n$ in $G$ is contained in a perfect matching of $G$.

Bicritical graphs play a central role in the decomposition theory of graphs in terms of their maximum matchings. In [10], it is proved that every $n$-extendable graph is also $(n - 1)$-extendable and also that any 2-extendable graph is either 1-extendable bipartite or bicritical. Thus for non-bipartite graphs, if we let $G_k$ denote the class of those which are $k$-extendable and $G_b$ those which are bicritical, we have a nested sequence of graph classes: $G_1 \supset G_b \supset G_2 \supset G_3 \ldots$. Moreover, it is known that each of the containments is proper. In [7], Lovsz and Plummer point out that the structure of bicritical graphs is far from completely understood. We show a necessary and sufficient condition to characterize all bicritical graphs in this paper.

The concept of bicritical graphs is also generalized to $n$-critical graphs and there has been some research on this topic (see [2]-[6]). In [12], Yu gives a Tutte style necessary and sufficient condition for $n$-critical graphs. However, it does not help to design an efficient algorithm to determine the $n$-critical graphs. Using our necessary and sufficient condition, we can design a highly efficient algorithm to determine all bicritical graphs.

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2 A Necessary and Sufficient Condition for Bicritical Graphs

In this section, we give a necessary and sufficient condition for bicritical graphs.

Theorem 1. Let $G$ be a graph with a perfect matching $M_0$, then the following propositions are equivalent:

1. $G$ is bicritical;
2. For any perfect matching $M$ and any two different vertices $x$ and $y$ in $G$, there is an $M$-alternating path between $x$ and $y$, which starts and ends with edges in $E(G) \setminus M$.
3. For every pair of vertices $x$ and $y$ in $G$, there is an $M_0$-alternating path between $x$ and $y$, which starts and ends with edges in $E(G) \setminus M_0$.

Proof. We show that (1) $\Rightarrow$ (2) $\Rightarrow$ (3) $\Rightarrow$ (1).

1. (1) $\Rightarrow$ (2): Suppose that $G$ is bicritical, then for any perfect matching $M$ in $G$ and any two vertices $u, v \in V(G)$, $G - \{u, v\}$ has a perfect matching. For any pair of vertices $x, y \in V(G)$, we discuss in two cases:
   
   Case 1. $xy \notin M$. Suppose that $xu, yv \in M$. $G - \{u, v\}$ has a perfect matching. Let $M_1 = M - \{xu, yv\}$ be a matching in $G - \{u, v\}$, then both $x$ and $y$ are $M_1$-unsaturated. So, there must be an $M_1$-augmenting path between $x$ and $y$, otherwise, according to Berge’s theorem, $M_1$ is the maximum matching in $G - \{u, v\}$, which contradicts the fact that $G$ is bicritical.

   Case 2. $xy \in M$. For $yw \in E(G) \setminus M$, since $G$ is 1-extendable, $yw$ is contained in a perfect matching $M'$. Let $G_1 = G[M \Delta M']$, where $M \Delta M'$ denotes the symmetric difference of $M$ and $M'$. Then the component of $G_1$ that contains $xy$ is an even cycle with edges alternatively in $M$ and $M'$, and we have an $M$-augmenting path between $x$ and $y$ as desired. (see Figure 2).

2. (2) $\Rightarrow$ (3) follows trivially since (3) is a special case of (2).

3. (3) $\Rightarrow$ (1) For any two vertices $u, v \in V(G)$, if $uv \in M_0$, then $G - \{u, v\}$ has a perfect matching $M_0 \setminus \{uv\}$. Suppose $uv \notin M_0$, and $xu, yv \in M_0$. By (3), there is an $M_0$-alternating path $P$ between vertices $x$ and $y$, which starts and ends with edges in $E(G) \setminus M_0$, and $P$ is an $M_0 \setminus \{xu, yv\}$-augmenting path in $G - \{u, v\}$. By exchanging edges in $M_0$ and edges not in $M_0$ along $P$, we have a perfect matching in $G - \{u, v\}$. Thus, $G$ is bicritical.

We call such an $M$-alternating path $P$ between two $M$-saturated vertices $x$ and $y$, which starts and ends with edges in $E(G) \setminus M$ a pseudo-$M$-alternating path.

Fig. 1. $G[M \Delta M']$, $M$ broken lines, $M'$ solid lines.
3 Description of the Algorithm

First, we can use an $O(\sqrt{n}|E|)$ algorithm at hand to find a perfect matching $M_0$ in a given graph $G$ (see [8]), if such $M_0$ cannot be found, then $G$ is not bicritical, and we can make our conclusion at once. Otherwise, for any two vertices, say, $x$ and $y$, we try to find an $M_0$-alternating path between them in $G$, which starts and ends with edges in $E(G) \setminus M_0$. This should be done for all pairs of vertices, and if there exists such a path between every pair of vertices in $G$, then according to Theorem $\text{[1]}$, $G$ is bicritical, otherwise, $G$ is not bicritical. To determine bicritical graphs, we develop an algorithm, MAPS($x$, exist) in the main recognition program to find all the $M_0$-alternating paths which start and end with edges in $E(G) \setminus M_0$ in $G$ between a given vertex, say $x$ and every other vertex in $G$ in $O(|E|)$ time.

The main recognition program to determine bicritical graphs

1. exist:=true;       // $O(1)$ time
2. Construct an adjacency list for $G$;    // $O(|E|)$ time
3. Try to find a perfect matching $M_0$ in $G$;     // $O(\sqrt{n}|E|)$ time
4. If such $M_0$ does not exist then exit; ($G$ is not bicritical)
5. else for each of $|V| - 1$ vertices in $G$, say $x$ do
   begin
6. MAPS($x$, exist);  // pseudo-$M_0$-alternating path search procedure
7. If exist=false then exit ($G$ is not bicritical')     // $O(1)$ time
   end
8. Exit ($G$ is bicritical)                               // $O(1)$ time

Our algorithm MAPS procedure is derived from the idea inspired by Edmond, there are some differences though, mainly in the handling of blossoms and in the way we extend an $M_0$-alternating tree. We use an adjacency list to store $G$. Suppose that all vertices have indices in sequence, and for all vertices, there exists a table in the storage, which is of the form below.

<table>
<thead>
<tr>
<th>Vertex ($i$)</th>
<th>1</th>
<th>2</th>
<th>3</th>
<th>...</th>
<th>$n$</th>
</tr>
</thead>
<tbody>
<tr>
<td>$m$</td>
<td></td>
<td></td>
<td></td>
<td></td>
<td></td>
</tr>
<tr>
<td>$f$</td>
<td></td>
<td></td>
<td></td>
<td></td>
<td></td>
</tr>
<tr>
<td>$b$</td>
<td></td>
<td></td>
<td></td>
<td></td>
<td></td>
</tr>
<tr>
<td>$p$</td>
<td></td>
<td></td>
<td></td>
<td></td>
<td></td>
</tr>
</tbody>
</table>

In this table, for any given vertex $i$, $m(i)$ denotes the other vertex adjacent to $i$ with an edge in $M_0$, $f(i)$ represents the father of $i$, that is, in the construction of the alternating tree, the predecessor of $i$, and $b(i)$ is the root of the blossom that contains $i$. Finally, if there is a pseudo-$M_0$-alternating path between $x$ and $i$ when constructing an $M_0$-alternating tree rooted with $x$, then $p(i) = 1$, otherwise $p(i) = 0$. Initially, let $f(i) := 0$; $b(i) := i$; $p(i) = 0$ for all $i$.

Construct the adjacency list and the table for $G$, which needs $O(|E|)$ time.

An alternating tree with root $x$ with respect to a matching $M$ is a subgraph of $G$, $\text{Tree} = (V,T)$, which is connected and without a cycle, such that:
1. $x \in v$ is the root of the tree, which is $M$-unsaturated;
2. for each $i \in V$, the unique path $L_T(i)$ of the tree between $x$ and $i$ is an alternating path.

The alternating tree of Figure 3 illustrates this definition.

![Fig. 2. An alternating tree with root $x = 1$. Heavy edges represent matching edges](image)

We shall say that the vertices of $G$ which belong to $V$ are labelled. The vertices of $V(G) - V$ are called unlabelled. Let $i \in V$ be a labelled vertex.

- If the path $L_T(i)$ has an even number of edges, then we say that the vertex $i$ is even. In particular, $x$, the root of the tree is even.
- If the path $L_T(i)$ has an odd number of edges, then the vertex $i$ is said to be odd. In an alternating tree, the odd vertices must all be of degree 2: this is precisely because there is always exactly one thin and one heavy edge incident to such a vertex. What’s more, it’s crucial to point out that if vertex $i$ is labelled odd, then there must exist a pseudo-$M_0$-alternating path between $x$ and $i$.
- If one walks through an alternating tree from a vertex $i$ to the root $x$, then the vertices which one meets are alternatively even and odd.

Now, we are ready to introduce our pseudo-$M_0$-alternating path search procedure (MAPS) between $x$ and every other vertex in $G$. We use breadth-first-search method (BFS) to handle every vertex in $G$. In each MAPS procedure, we need an even vertices queue $V_0$ to store even vertices. Denote by $V_0 \subset V(G)$ the subset of vertices labelled even. The construction of the alternating tree will proceed step by step, starting from

$$V = V_0 = \{x\}, T = \emptyset$$

At each stage, Tree $= (V,T)$ induces an alternating tree on $G$. The MAPS procedure is as follows:

The pseudo-$M_0$-alternating path search procedure (MAPS) between $x$ and every other vertex in $G$.

1. Set the even vertices queue $V_0$ empty;
2. Store $x$ in even vertices queue $V_0$, label $x$ even and set $p(x) := 1$;
3. Remove the first vertex $i$ from $V_0$;
4. Use BFS to handle $i$, that is, for every edge $e = (i, j)$ which is incident to $i$ and $j$ is not labelled odd, two cases are possible:
5. **Case 1:** $j$ is unlabelled;
   - Label $j$ odd;
   - Set $p(j) := 1$;
   - Get $m(j)$ from the table;
   - If $m(j)$ is unlabelled then
     - Add $m(j)$ to the even vertices queue $V_0$;
     - Write $f(j) := i, f(m(j)) := j$ to the table;
     - Label $m(j)$ even;
   
   // $T ← T + \{e\} + \{jm(j)\}$
   // $V ← V + \{j\} + \{m(j)\}$
   // $V_0 ← V_0 + \{m(j)\}$
6. **Case 2:** $j$ is also labelled even ($j \in V_0$)
   - $i' := i, j' := j$;
7. **While** $i' \neq b(i')$ or $j' \neq b(j')$ **do**
   - $i' := b(i')$;
   - $j' := b(j')$;
   - If $i' = j'$ then we do nothing **else**
     - $k := i', l := j'$;
8. **While** (neither $i'$ nor $j'$ has been marked) and ($i' \neq j'$) **do**
   - **Marked** $i'$ and $j'$
     - $i' := b(f(i'))$;
     - $j' := b(f(j'))$;
   - **end**
   - If $j'$ has been marked then $r := j'$
   - **else** $r := i'$
   **end**

*Note 1.* $r$, which has at least degree 3 (unless $r = x$), must be labelled even. The subpath $L_i$ and $L_j$ of $L_T(i)$ and $L_T(j)$, on the one hand between $r$ and $i$, and on the other hand between $r$ and $j$, are even and $\mu = L_i + L_j + \{e\}$ is certainly of odd cardinality. The matching $M_\mu = M_0 \cap \mu$ is a maximum matching of $\mu$ and $r$, the only $M_\mu$-unsaturated vertex of the cycle, is linked with $x$ by an even alternating path $L_T(r)$. Such a cycle $\mu$ is called a blossom of $M_0$, and the vertex $r$ is called the root of the blossom. When a blossom has been found, each vertex in it must have a pseudo-$M_0$-alternating path that links it to the root $x$. Now we use a supplementary queue $V_1$ to store the vertices of $\mu$ that are labelled odd and notice that we need not make any changes to those vertices that have already been included in some other blossoms. This can be done as follows.
9. While \( k \neq r \) do 
   begin 
   If \( k \) is labelled odd then 
     begin 
     Put \( k \) into \( V_1 \); 
     Re-label \( k \) even 
     end 
   else \( p(k) := 1; \) 
   \( b(k) := r \) 
   Write \( b(k) \) to the table; 
   \( k := f(k) \) 
   end 
10. While \( l \neq r \) do 
   begin 
   If \( l \) is labelled odd then 
     begin 
     Put \( l \) into \( V_1 \); 
     Re-label \( l \) even 
     end 
   else \( p(l) := 1 \) 
   \( b(l) := r \) 
   Write \( b(l) \) to the table 
   \( l := f(l) \) 
   end 
11. Add \( V_1 \) to the tail of \( V_0 \); 
end of case 2; 
12. If \( V_0 \) is not empty then go to 4 
13. for every vertex \( i \) of \( G \) do 
   if \( p(i) = 0 \) then exit(exist:=false); 
14. exit(exist:=true).

4 The Correctness and the Time Complexity of the Program

Lemma 1. If there exist pseudo-\( M_0 \)-alternating paths between \( x \) and every other vertex \( y \) in \( G \), then MAPS can find them.

Proof. If \( y \) is labelled odd then the lemma trivially follows. If \( y \) is labelled even then only when it is contained in a blossom, there is a pseudo-\( M_0 \)-alternating path between \( x \) and \( y \). The last case is that the \( M_0 \)-alternating tree rooted with \( x \) fails without reaching \( y \). In this case, we prove that there is no pseudo-\( M_0 \)-alternating path between \( x \) and \( y \) in \( G \) as follows:

Let \( u, v \) be the vertices that are adjacent to \( x \) and \( y \) respectively, with matching edges. If there exists an \( M_0 \)-alternating path between \( x \) and \( y \) in \( G - \{u, v\} \), then it must be an \( M_0 \)-augmenting path. When the alternating tree with
root \( x \) fails to reach \( y \), the matching is maximum (with only \( x \) and \( y \) \( M_0 \)-unsaturated). Thus, according to Berge’s theorem, there is no \( M_0 \)-augmenting path in \( G - \{u, v\} \). So the lemma follows.

**Theorem 2.** The program that determines whether any given simple graph \( G \) is bicritical is correct.

**Proof.** For every pair of vertices \( x \) and \( y \) in \( G \), the program is to check whether there exits a pseudo-\( M_0 \)-alternating path between \( x \) and \( y \), which starts and ends with edges in \( E(G) \setminus M_0 \). Thus, the theorem is a direct result from Theorem 1 and Lemma 1.

**Theorem 3.** In the worst case, the time complexity of the program is bounded by \( O(|V||E|) \).

**Proof.** In the following proof, denote \( |V| \) by \( N \). First, let’s look at the main recognition program to determine bicritical graphs: obviously, program sections 5–7 are a loop that checks each vertex in \( G \), in the worst case, we must construct \( N - 1 \) alternating trees. The problem is therefore that of determining the number of elementary operations (additions, comparisons, and access to stored space) needed to construct an \( M_0 \)-alternating tree, that is, the time complexity of the procedure MAPS.

The structure of the graph \( G \) is supposed to be described by an adjacency list giving, for each vertex \( i \), those vertices which are adjacent to \( i \). Each edge \( e = (i, j) \) appears thus twice in the list.

In the procedure MAPS, program sections 1–3 need \( O(1) \) time. Each edge is examined at most twice in our BFS in accordance to their appearance in the adjacency list. So the number of operations that work on the adjacency list is proportional to \( |E| \).

When a blossom \( \mu_i \) has been detected, the number of operations necessary to change the labels and the updates \( b(i) \) of the vertices of the blossom is proportional to \( |\mu_i| \). Notice that the number of blossoms which might appear in the course of the construction of an alternating tree is at most \( \lfloor N/2 \rfloor \) (the integer part of \( N/2 \)). Since every vertex is handled at most once in a blossom, the total sum of \( |\mu_i| \), \( i \) from 1 to \( \nu \), is at most \( N + N/2 \) (\( N/2 \) is the maximal number of the roots of all blossoms). Thus, the number of all operations of this type is proportional to \( N \).

For those edges that won’t form blossoms when they are added to the alternating tree, the time complexity over all is \( O(k) \), \( k \) is the number of such edges, and for those edges that will form blossoms when they are added to the alternating tree, the time complexity over all is \( O(N + N/2) \), thus in the procedure MAPS from program sections 4–12, we need at most \( O(k + N + N/2) \) time, that is, \( O(|E|) \) time to construct an \( M_0 \)-alternating tree, Section 13 needs \( O(N) \) time and section 14 needs \( O(1) \) time, so our MAPS needs \( O(|E|) \) time, and for all \( N \) alternating trees, we need \( O(N|E|) \). Thus the theorem follows.
5 Remarks

If we use the algorithm derived from the definition of bicritical graphs, that is, choose every pair of vertices $x$ and $y$ of $G$, delete $x$ and $y$ together with edges incident to them, and then check whether $G - \{u, v\}$ has a perfect matching, then the time complexity of such an algorithm will come out as $O\left(N^{5/2} |E|\right)$, which is obviously inferior to the algorithm we get above, with the space complexity almost the same. So, our algorithm has higher efficiency.

References

Layered Drawings of Graphs with Crossing Constraints

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Abstract. We study the problem of producing hierarchical drawings of layered graphs when some pairs of edges are not allowed to cross. We show that deciding on the existence of a drawing satisfying at least \( k \) constraints from a given set of non-crossing constraints is NP-complete even if the graph is 2-layered and even when the permutation of the vertices on one side of the bipartition is fixed. We also propose simple constant-ratio approximation algorithms for the optimization version of the problem and we discuss how to extend the well-known hierarchical approach for creating layered drawings of directed graphs with the capability of minimizing the number of edge crossings while maximizing the number of satisfied non-crossing constraints.

1 Introduction

The problem of embedding graphs in the plane when only some pairs of edges are allowed to cross, known in literature as realizability, has been formally introduced in [12]. It finds application, e.g., in VLSI layout, where the crossings between certain pairs of edges must be avoided due to the physical realization of connectors, and in constrained graph layout, where users can specify requirements on the visualization that should be fulfilled by the drawing algorithm. In these settings, it is fundamental to decide if an embedding of a graph satisfying a given set of non-crossing constraints exists, and, if the answer is positive, to find it. Realizability has been proved to be NP-hard in [10], but surprisingly it is not known to belong to the class NP, turning out to be a very interesting problem from a theoretical perspective. Its theoretical relevance is mostly due to the connection with the classical themes of planarity and crossing numbers and with the recognition of string graphs, i.e., intersection graphs of curves in the plane. The relation between realizability and string graphs is addressed in [10,11]; here we limit to point out that no finite algorithm for the recognition of string graphs is known and that, given a graph \( G \), it is possible to build an instance of realizability which can be satisfied if and only if \( G \) is a string graph [11]. It is also worth

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observing that realizability is a generalization of the well-known planarity problem, which requires to decide if a graph is planar and clearly coincides with the special case where no pair of edges is allowed to cross. We recall that planarity can be checked in linear time [9].

This paper is concerned with the realizability problem in the special case of hierarchical drawings [3], i.e., drawings where vertices are constrained to lie on a set of parallel lines and edges are represented as polygonal chains. The hierarchical drawing convention is widely used to represent procedure call dependencies, class hierarchies, is-a relationships, and several economic and social structures [3, 5, 14]. In the remainder of the paper we will refer to the vertex set on each line as layer or level, and sometimes to the drawing itself as layered drawing. Without loss of generality we also assume to deal with proper layered graphs, i.e., layered graphs whose edges span only consecutive levels: it is always possible to reduce to this case by properly adding dummy vertices to split edges with end-points in non-consecutive layers. Under this hypothesis each edge in a hierarchical drawing is simply represented by a straight-line. The main question that we address is the following: Given a proper layered graph $G$, a set of edge pairs $C$, and an integer $k \geq 0$, does a hierarchical drawing of $G$ exist s.t. at least $k$ edge pairs in $C$ do not cross? We refer to this problem as hierarchical realizability (HR) and to the elements of $C$ as non-crossing constraints. We study the computational complexity of some variants of hierarchical realizability and we describe approximation strategies for the optimization versions of these problems, that require to find a maximum realizable subset of $C$. We also show how algorithms for the variants that we consider can be used as subroutines to boost existing graph drawing algorithms to support non-crossing constraints. Our results can be easily generalized to constraints concerning the relative positions of vertices within each layer, though for simplicity we will not consider them in this paper.

In more detail, the presentation of our results is organized as follows. In Section 2 we prove that deciding on the realizability in case of hierarchical drawings is not easier than deciding on the realizability for generic drawings: we show that HR, which clearly belongs to $NP$, remains NP-hard even if $G$ is 2-layered and $k = |C|$. It is interesting to observe that the strictly related problem of deciding if a layered graph has a hierarchical drawing without edge crossings can be solved in linear time [8]. From an optimization point of view, it is trivial to approximate a maximum cardinality subset of realizable constraints with expected ratio 2, yet we can exhibit instances for which this bound is tight. In addition to the general case, in Section 3 we consider a one-sided realizability problem where $G$ is 2-layered and the vertex ordering of a layer is fixed; we show this problem to be NP-hard for a generic $k$, approximable with constant ratio 2, and polynomial if $k = |C|$. Based on the one-sided formulation, in Section 4 we finally address the constrained edge crossing minimization problem in layered drawings, i.e., the problem of drawing hierarchically a graph with the minimum number of edge crossing while respecting a set of non-crossing constraints. We consider the well-known Sugiyama’s approach for producing layered drawings of directed
graphs [14] and we extend it in order to deal with non-crossing constraints. In particular, we provide a mechanism for identifying a subset $C' \subseteq C$ of realizable constraints, and then we show how to minimize edge crossings while satisfying all the constraints in $C'$.

2 Hierarchical Realizability

In this section we prove that deciding about the hierarchical realizability of a proper layered graph under a given set of non-crossing constraints is NP-complete. From an optimization point of view, we show that a maximum cardinality subset of realizable constraints can be easily approximated with expected ratio 2 and we exhibit instances for which this bound is tight.

Lemma 1. $HR \in NP$.

Proof. Since edge crossings in layered drawings depend only on the permutation of the vertices within each layer, a non-deterministic algorithm has simply to guess an ordering of the vertices of each layer and to check in polynomial time the realization of at least $k$ constraints.

In the following we show the NP-hardness of HR on two-layered (i.e., bipartite) graphs. For brevity, we call bipartite straight-line drawing a hierarchical drawing with only two layers. Let $B = (V_0, V_1, E)$ be a bipartite graph with $n = n_0 + n_1$ vertices and $m$ edges. Without loss of generality, we assume that for any $(a, b) \in E$ $a \in V_0$ and $b \in V_1$. We also denote by $\pi_0$ and $\pi_1$ two permutations of the vertices in $V_0$ and $V_1$, respectively. Let $C \subseteq \binom{E}{2}$ be a set of pairs of edges. The hierarchical realizability problem on bipartite graphs consists of asking for a set $C' \subseteq C$, $|C'| \geq k$, and for a bipartite straight-line drawing of $B$ such that for any $(a, b; c, d) \in C'$ edges $(a, b)$ and $(c, d)$ do not cross in the drawing. More formally, the problem can be stated as follows:

Bipartite realizability (in short, BR): Let $B = (V_0, V_1, E)$ be a bipartite graph, let $C$ be a set of non-crossing constraints, and let $k$ be a positive integer. Do two permutations $\pi_0$ and $\pi_1$ exist such that the bipartite straight-line drawing of $B$ induced by $\pi_0$ and $\pi_1$ satisfies at least $k$ constraints?

In the special case where the permutation of the vertices in a layer (e.g., $V_0$) is fixed, we have the one-sided bipartite realizability problem (in short, OBR), whose complexity and approximability are discussed in Section 3.

Before proving the NP-completeness of BR, we state it in a slightly different form which is more convenient for our purposes. Indeed, it is easy to see that asking for two permutations satisfying a set of non-crossing constraints $C'$ is equivalent to asking for two bijective functions $\pi_0$ and $\pi_1$ such that:

$$\forall(a, x; b, y) \in C' \quad (\pi_0(a) - \pi_0(b)) \cdot (\pi_1(x) - \pi_1(y)) \geq 0 \quad (1)$$

Based on this observation, we prove the NP-hardness of BR by means of a polynomial time reduction from a total ordering problem arising in the computational biology field, called betweenness, which has been shown to be NP-complete in [13] and is defined as follows:
**Betweenness:** Given a finite set \( S \) of real variables and a set of triples \( T \subseteq S^3 \), does an ordering of the variables exist such that for each \((a, b, c) \in T\) variable \( b \) is positioned between variables \( a \) and \( c \) in the ordering?

We refer to the triples in \( T \) as *ordering constraints*. A solution for betweenness is a bijective function \( f : S \rightarrow \{1..|S|\} \) such that:

\[
\forall(a, b, c) \in T \quad (f(a) - f(b)) \cdot (f(b) - f(c)) \geq 0
\]

We remark that, in spite of the apparent similarity of Equation 1 and Equation 2, the former involves four distinct vertices and two different functions, while the latter only three variables and a single function. A one-to-one correspondence is therefore not straightforward.

**Lemma 2.** Bipartite realizability is NP-hard.

**Proof.** Moving from an instance of betweenness, we build an instance of BR as follows. The bipartite graph \( B = (V_0, V_1, E) \) is such that: (a) \( V_0 = S \); (b) \( V_1 \) contains a distinct vertex \( x' \) for each item \( x \in S \), i.e., \( V_1 = \{x' : x \in S\} \); (c) \( E = \{(x, x') : x \in S\} \cup \{(a, b'), (b, c') : (a, b, c) \in T\} \). The set of non-crossing constraints \( C \) is the union of two disjoint sets \( C_1 \) and \( C_2 \), where \( C_1 = \{(x, x'; y, y') : x, y \in S, x \neq y\} \) and \( C_2 = \{(a, b'; b, c') : (a, b, c) \in T\} \). Note \(|C_1| = |S|(|S| - 1)\) and \(|C_2| = |T| \cdot k = |C_1| + |C_2|\), i.e., the problem is to decide if all the non-crossing constraints can be realized.

It is easy to verify that the reduction above requires \( O(|T| + |S|^2) \) time. To prove that \( k \) constraints in \( C \) can be realized iff the instance of betweenness admits a solution we exploit the one-to-one correspondence between ordering constraints in \( T \) and crossing constraints in \( C_2 \) and the fact that realizing all the constraints in \( C_1 \) forces \( \pi_0(a) = \pi_1(a') \) for any pair of corresponding vertices \( a \in V_0 \) and \( a' \in V_1 \).

Let us first suppose that two permutations \( \pi_0 \) and \( \pi_1 \) exist able to satisfy all the constraints in \( C \). Then \( f = \pi_0 \) represents a solution for betweenness. Indeed, let \((a, b, c) \in T\) be an ordering constraint in the instance of betweenness. Let us consider the constraints \((b, b'; c, c') \in C_1 \) and \((a, b'; b, c') \in C_2\); these constraints necessarily exist by construction. Since \( \pi_0 \) and \( \pi_1 \) satisfy all the non-crossing constraints, by Equation 1 we know that: \((\pi_0(a) - \pi_0(b)) \cdot (\pi_1(b') - \pi_1(c')) \geq 0\) and \((\pi_0(b) - \pi_0(c)) \cdot (\pi_1(b') - \pi_1(c')) \geq 0\). By multiplying the left members of these inequalities, we obtain \((\pi_0(a) - \pi_0(b)) \cdot (\pi_0(b) - \pi_0(c)) \geq 0\), proving that the ordering constraint \((a, b, c) \in T\) is satisfied by \( f = \pi_0 \).

Conversely, let us assume that the instance of betweenness admits a solution \( f \) and let us define \( \pi_0 = f \) and \( \forall a' \in V_1 \pi_1(a') = f(a) \). Note that \( \forall a \in S \pi_0(a) = \pi_1(a') \). Any constraint \((a, a'; b, b') \in C_1 \) is trivially satisfied by \( \pi_0 \) and \( \pi_1 \), since \( \pi_0(a) - \pi_0(b) = \pi_1(a') - \pi_1(b') \) and therefore their product is \( \geq 0 \). Let \((a, b'; b, c') \) be any constraint in \( C_2 \). As above, \((\pi_0(b) - \pi_0(c)) \cdot (\pi_1(b') - \pi_1(c')) \geq 0\). Moreover, since \( \pi_0 = f \) satisfies constraint \((a, b, c) \in T\), by Equation 2 \((\pi_0(a) - \pi_0(b)) \cdot (\pi_0(b) - \pi_0(c)) \geq 0\). By multiplying the left members of the two inequalities we finally get \((\pi_0(a) - \pi_0(b)) \cdot (\pi_1(b') - \pi_1(c')) \geq 0\), i.e., constraint \((a, b'; b, c') \) is satisfied by \( \pi_0 \) and \( \pi_1 \), as well.
The NP-completeness of HR immediately follows from Lemma 1 and Lemma 2.

**Theorem 1.** Hierarchical realizability is NP-complete.

We now consider the problem of approximating a maximum set of realizable constraints. If we choose, for each layer of $G$, a random ordering of the vertices in that layer, we approximate a maximum set of realizable constraints (chosen from $C$) with expected ratio $2$.

If $c = |C|$ and $c_i = (a, x; b, y)$, $1 \leq i \leq c$, is a non-crossing constraint involving vertices from any two consecutive layers $L_j$ and $L_{j+1}$, four orderings among $a, b \in L_j$ and $x, y \in L_{j+1}$ are possible, yet only two of them realize $c_i$. Hence, the probability $p_i$ of having constraint $c_i$ satisfied is equal to $\frac{1}{2}$. Let $X_i$ be a random variable equal to $1$ if constraint $c_i$ is realized, $0$ otherwise. The quantity $\sum_{i=1}^{c} X_i$ represents the number of realized constraints and the expected value for it is then equal to $E[\sum_{i=1}^{c} X_i] = \sum_{i=1}^{c} E[X_i] = \sum_{i=1}^{c} p_i = \frac{c}{2}$.

We now show that the ratio above is tight, i.e., no algorithm can do better than trying to satisfy half of the constraints in $C$, unless being able to recognize non-realizable instances of HR. The following lemma is useful at this aim:

**Lemma 3.** The number of non-crossing pairs of non-incident edges in any bipartite straight-line drawing of $K_{n,n}$ is equal to $\left(\frac{n(n-1)}{2}\right)^2$.

**Proof.** Any bipartite straight-line drawing of $K_{n,n}$ has the same number $x$ of non-crossing pairs of non-incident edges. Then:

$$x = \frac{1}{2} \sum_{i=1}^{n} \sum_{j=1}^{n} \left( \sum_{k<i} \sum_{h<j} 1 + \sum_{k>i} \sum_{h>j} 1 \right) = \left(\frac{n(n-1)}{2}\right)^2$$

Let $G = K_{n,n}$ and let $C$ contain a constraint for any pair of non-incident edges. It is easy to see that $c = |C| = \frac{n^2(n-1)^2}{2}$ and therefore, due to Lemma 3, the number of realized constraints is always equal to $\frac{c}{2}$ independently of the permutation of the two layers.

3 One-Sided Bipartite Realizability

In this section we focus on the one-sided bipartite realizability problem, that turns out to be very useful in the design of algorithms for solving HR, as we will show in Section 4. Unfortunately, we can prove that it is NP-complete, yet easier than BR, being polynomially solvable if we are interested in the existence of a solution satisfying all the constraints in $C$ (i.e., for $k = |C|$).

**NP-completeness.** For simplicity we prove the NP-hardness of one-sided bipartite realizability in two steps, making use of a simple generalization of this problem which takes into account also crossing constraints in addition to the non-crossing ones. We refer to the generalized problem as one-sided bipartite realizability with mixed constraints (in short, MBR). MBR is stated exactly as OBR, except for taking into account a set of crossing constraints $I \subseteq \binom{E}{2}$ with
the following meaning: for any \((a, b; c, d) \in I\), edges \((a, b)\) and \((c, d)\) should cross in the bipartite drawing of graph \(B\).

With arguments similar to those used in the proof of Lemma 1, it is easy to see that both OBR and MBR belong to NP. We now prove the NP-hardness of MBR by means of a polynomial-time reduction from the acyclic subgraph problem. Then we further reduce MBR to our original realizability problem OBR. We recall that the acyclic subgraph problem is stated as follows:

\textit{Acyclic subgraph} (in short, AS): given a directed graph \(G = (V, A)\) and a positive integer \(h\), does a set \(A' \subseteq A\) exist such that \(|A'| \geq h\) and the graph \(G' = (V, A')\) is acyclic?

This problem has been proved to be NP-complete in \cite{[6]}.

\textbf{Lemma 4.} \textit{One-sided bipartite realizability with mixed constraints is NP-hard.}

\textit{Proof.} Given an instance of AS, we build an instance of MBR as follows. The bipartite graph \(B = (V_0, V_1, E)\) is such that: (a) \(V_1 = V\); (b) \(V_0\) contains a distinct vertex \(x'\) for each vertex \(x \in V\), i.e., \(V_0 = \{x' \text{ s.t. } x \in V\}\); (c) \(\forall (x, y) \in A\) both \((x', y)\) and \((y', x)\) belong to \(E\). Permutation \(\pi_0\) is any ordering of the vertices in \(V_0\). For each arc \((x, y) \in A\) we set up a constraint \((x', y; y', x)\): if \(\pi_0(x') < \pi_0(y')\) we add \((x', y; y', x)\) to \(I\), otherwise we add \((x', y; y', x)\) to \(C\). \(k = h\). Figure 1 provides an example of this reduction. A directed graph \(G\) is given in Figure 1(a) and the bipartite graph \(B\) obtained from it is shown in Figure 1(b), together with the whole set of constraints obtained choosing \(\pi_0 = \langle d', a', c', b' \rangle\).

\begin{figure}[h]
\centering
\includegraphics[width=\textwidth]{example.png}
\caption{From acyclic subgraph to one-sided realizability with mixed constraints.}
\end{figure}

It is worth observing that there is a one-to-one correspondence between constraints and arcs of the digraph and that the constraint \((x', y; y', x)\) associated to arc \((x, y)\) belongs to \(C\) if and only if \(\pi_0(x') > \pi_0(y')\). To prove the correspondence between solutions of AS and solutions of MBR we now exploit a basic idea: any permutation of the vertices of a digraph partitions its arcs into left-to-right and right-to-left arcs. We recall that a topological sort of an acyclic digraph yields no right-to-left arc. Based on this idea, Figure 2 points out the meaning of both crossing and non-crossing constraints: if arc \((x', y')\) is left-to-right according to \(\pi_0\), then we force \(x\) and \(y\) to have the same ordering as \(x'\) and \(y'\) using a crossing constraint (see Figure 2(a)); if arc \((x', y')\) is right-to-left according to \(\pi_0\), then we force \(x\) and \(y\) to have the opposite ordering of \(x'\) and \(y'\) using a non-crossing constraint (see Figure 2(b)).
Fig. 2. Crossing and non-crossing constraints: (a) \((x', y; y', x) \in I\); (b) \((x', y; y', x) \in C\). In both cases arc \((x, y)\) is forced to be a left-to-right arc.

Roughly speaking, our constraints force the presence of as many left-to-right arcs as possible. Proving that any permutation \(\pi_1\) realizes \(k\) constraints from \(I \cup C\) iff it yields an acyclic subgraph of \(G\) containing \(h = k\) arcs is now easy. Let \(c = (x', y; y', x)\) be a constraint from \(I \cup C\). If \(c \in I\) (i.e., \(\pi_0(x') < \pi_0(y')\)), then \(c\) is realized by \(\pi_1\) if and only if \(\pi_1(x) < \pi_1(y)\). Analogously, if \(c \in C\) (i.e., \(\pi_0(x') > \pi_0(y')\)), then \(c\) is realized by \(\pi_1\) if and only if \(\pi_1(x) < \pi_1(y)\). In both cases the satisfaction of constraint \(c\) implies that arc \((x, y)\) is a left-to-right arc according to \(\pi_1\), and vice-versa. Therefore, any permutation \(\pi_1\) satisfies a number of constraints equal to the number of left-to-right arcs that it induces.

In the example in Figure 1, choosing e.g. \(\pi_1 = \langle a, b, c, d \rangle\) partitions the arcs of \(G\) as in Figure 1(c) and yields the bipartite straight-line drawing of \(B\) shown in Figure 1(d). In this drawing only constraint \((d', a; a', d) \in I\) is not satisfied; note this constraint corresponds to the unique dotted arc \((d, a) \in A \setminus A'\).

**Theorem 2.** One-sided bipartite realizability is NP-hard.

**Proof.** Let us consider an instance of MBR specified by a bipartite graph \(B = (V_0, V_1, E)\), a set of crossing and non-crossing constraints \(I \cup C\), a permutation \(\pi_0\), and an integer \(k\) (w.l.o.g. in the following we assume that \(\pi_0(a) < \pi_0(b)\) for any constraint \((a, x; b, y) \in I \cup C\)). We build a corresponding instance of OBR specified by a bipartite graph \(B' = (V'_0, V'_1, E')\), a set of non-crossing constraints \(C'\), a permutation \(\pi'_0\), and an integer \(k'\) as follows.

For each crossing constraint \((a, x; b, y) \in I\) we create a vertex named \(\hat{ab}\), an edge \((\hat{ab}, y)\), and a non-crossing constraint \((\hat{ab}, y; a, x)\). We denote the sets of all the new vertices, the new edges, and the new constraints with \(\hat{V}_0\), \(\hat{E}\), and \(\hat{C}\), respectively. We then set \(V'_0 = V_0 \cup \hat{V}_0\), \(V'_1 = V_1\), \(E' = E \cup \hat{E}\), \(C' = C \cup \hat{C}\), and \(k' = k\). At last, we choose any permutation \(\pi'_0\) such that all vertices in \(\hat{V}_0\) are to the left of vertices in \(V_0\) (their relative positions are not important for our purposes), while the vertices in \(V_0\) retain the same ordering as in \(\pi_0\).

In the following we show that any permutation \(\pi_1\) satisfies the same number of constraints in the instance of MBR and in the corresponding instance of OBR, proving a one-to-one correspondence between the solutions of the two problems. It is obvious that any permutation \(\pi_1\) satisfies the same constraints from \(C\) in both instances. Moreover, there is a one-to-one correspondence between constraints in \(I\) and constraints in \(\hat{C}\): let \((a, x; b, y) \in I\) and let \((\hat{ab}, y; a, x) \in \hat{C}\) be the corresponding non-crossing constraint. Recall that by construction we
have \( \pi'_0(ab) < \pi'_0(a) < \pi'_0(b) \). Therefore, the crossing constraint \((a,x;b,y)\) is realized by \( \pi_1 \) if and only if the non-crossing constraint \((ab,y;a,x)\) is, proving that each satisfied constraint from \( I \) corresponds to a satisfied constraint from \( \hat{C} \) and vice-versa.

**Approximability and polynomiality.** In order to prove that OBR is polynomial if \( k = |C| \) and that a maximum set of realizable constraints can be approximated within a constant ratio we use an approximation preserving reduction \([1]\) to the maximum acyclic subgraph problem. This problem is the optimization version of AS: given a digraph \( H = (N,A,w) \) with positive arc weights \( w \), it requires to find a maximum weight set of arcs \( A' \subseteq A \) such that the subgraph \( H' = (N,A') \) is acyclic.

Let \( B = (V_0,V_1,E) \) and \( \pi_0 \) specify an instance of one-sided bipartite realizability. W.l.o.g. we assume that for each \((a,x;b,y) \in C\) it holds \( \pi_0(a) < \pi_0(b) \); otherwise we change \( C \) by removing \((a,x;b,y)\) and adding the analogous constraint \((b,y;a,x)\). In the reduction we build a weighted constraint digraph \( H = (N,A,w) \) as follows: \( N = V_1; \; A = \{(x,y) : \exists a,b \in V_0 \text{ s.t. } (a,x;b,y) \in C \}; \forall (x,y) \in A, w(x,y) \) is the number of constraints of \( C \) containing \( x \) and \( y \) in the second and fourth positions, respectively. In other words, the weight of arc \((x,y)\) is the number of constraints that require \( \pi_1(x) < \pi_1(y) \) in order to be realized.

An almost trivial observation is that \( H \) is acyclic iff all the constraints in \( C \) are realizable. This means that OBR can be solved in linear time if \( k = |C| \), thanks to the fact that both the previous reduction and checking the acyclicity of the constraint digraph can be accomplished in linear time.

The following property also holds: each \( A' \subseteq A \) such that the subgraph \( H' = (N,A') \) is acyclic corresponds to a realizable set of constraints \( C' \subseteq C \) with \( |C'| \geq w(A') \). Indeed, if \( \pi_1 \) is a topological sort of \( H' \), then \( \pi_1 \) satisfies at least all the constraints corresponding by construction to the arcs in \( A' \), whose quantity is \( w(A') = \sum_{(x,y) \in A'} w(x,y) \) (it could happen that \( \pi_1 \) satisfies more constraints if \( A' \) is not maximal). Finally, it is easy to see that the weight of a maximum acyclic subgraph of the constraint digraph equals the maximum number of realizable constraints, i.e., \( w(A^*) = |C^*| \).

Let \( r \geq 1 \) be any approximation ratio for the maximum acyclic subgraph problem. From the previous considerations we have:

\[
|C'| \geq w(A') \geq \frac{1}{r} \cdot w(A^*) = \frac{1}{r} \cdot |C^*|
\]

proving that \( C' \) is an \( r \)-approximate solution for one-sided bipartite realizability, as well. Hence, a maximum set of realizable constraints can be approximated with ratio 2 \([7]\).

### 4 Constrained Crossing Minimization

A widely used approach for creating layered drawings of directed graphs is presented in \([14]\) and is known in literature as **hierarchical approach.** In order to produce readable drawings, different aesthetic criteria are taken into account by
this method; among them the number of edge crossings deserves special attention. In this section we show how to extend the hierarchical approach in order to support non-crossing constraints. Our extension hinges upon the approximation algorithm for one-sided bipartite realizability presented in Section 3. We first briefly recall how Sugiyama’s algorithm works. Given a directed graph \(G\), it performs four main steps. (1) Cycle removal: the direction of some arcs is temporarily reversed to make \(G\) acyclic. (2) Layer assignment: each vertex is assigned with a level so that all the arcs “flow” in the same direction. Dummy vertices are added for arcs spanning more than two levels. (3) Crossing reduction: vertices in each level are ordered so as to minimize the total number of edge crossings. (4) Coordinates assignment: vertical coordinates are proportional to layers, horizontal coordinates depend on the permutation of the vertices within each layer. At this point the direction of the reversed arcs is also restored.

In order to extend the hierarchical approach to support non-crossing constraints we need to focus on steps 2 and 3, because neither the removal of cycles nor the assignment of coordinates are affected by this kind of constraints. In the following we denote with \(G\) the directed graph to be drawn and with \(C\) the set of non-crossing constraints to be realized in the drawing.

**Extending step 2.** We recall that in this step the directed graph \(G\) is transformed into a proper layered graph. The set of constraints \(C\) must be therefore changed accordingly, due to the fact that dummy vertices could be added to split an edge involved in some constraint. We let the algorithm assign layers to vertices as usual, where \(l(v)\) denotes the level of a vertex \(v\) after the layer assignment. Then we change \(C\) as follows.

Let \((a, x)\) and \((b, y)\) be any two arcs of \(G\) such that constraint \((a, x; b, y) \in C\); w.l.o.g. we can assume \(l(a) < l(x)\) and \(l(b) < l(y)\). If \(l(x) = l(a) + 1\) and \(l(y) = l(b) + 1\) we do not touch \(C\). Otherwise, at least a dummy vertex has been added to split \((a, x)\) or \((b, y)\) or both and constraint \((a, x; b, y)\) must be suitably replaced. Let \((u_k, u_{k+1}, \ldots, u_{k+s})\) be the path that replaced arc \((a, x)\), with \(u_k = a, u_{k+s} = x, k = l(a)\), and \(k + s = l(x)\). Analogously, let \((v_h, v_{h+1}, \ldots, v_{h+t})\) be the path that replaced arc \((b, y)\), with \(v_h = b, v_{h+t} = y, h = l(b), \) and \(h+t = l(y)\). Let also \([r, r+q] = [k, k+s] \cap [h, h+t]\). If \([r, r+q] \neq \emptyset\) we replace constraint \((a, x; b, y)\) with the constraints \((u_r, u_{r+1}; v_r, v_{r+1}) \ldots (u_{r+q-1}; u_{r+q-1}, v_{r+q})\). Otherwise we remove it, since it is going to be satisfied by any drawing obtained from layer assignment \(l\). Figure 3 shows an example of the two cases. It is easy to get convinced that an original constraint will be realized in the final drawing if and only if all the new constraints are satisfied.

Thanks to the preprocessing of set \(C\) described above, we can now assume to deal only with constraints involving edges that join vertices in consecutive layers. More formally, we assume that \(G\) has \(k\) layers, named \(L_0 \ldots L_k\), and that each \((a, x; b, y) \in C\) is such that \(\exists i \in [0..k-1]\) such that \(a, b \in L_i\) and \(x, y \in L_{i+1}\). Based on this assumption, we can partition \(C\) into \(k-1\) sets, named \(C_{0,1} \ldots C_{k-1,k}\), where \(C_{i,i+1}\) contains the constraints related to edges with endpoints in \(L_i\) and \(L_{i+1}\).
Extending step 3. Before attempting at minimizing edge crossings, we introduce a constraint realization step, aimed at finding a large set of realizable constraints. Then we devise a crossing minimization strategy able to take into account and satisfy the set of realizable constraints previously identified.

We recall that the most used strategy for crossing minimization in hierarchical drawings is based on a layer-by-layer sweep heuristic [3], that requires to repeatedly solve a crossing minimization problem on a bipartite graph with one side fixed. We therefore adapt the layer-by-layer sweep method to support non-crossing constraints, by repeatedly solving a one-sided bipartite realizability problem as follows.

A vertex ordering $\pi_0$ for $L_0$ is first randomly chosen. Then, for $i = 0$ to $k - 1$, the bipartite subgraph $B_i$ of $G$ induced by the vertices in $L_i \cup L_{i+1}$ and the set of constraints $C_{i,i+1}$ are considered. An approximated solution to OBR on this instance is found, keeping layer $L_i$ fixed. This leads to identify a set $C'_{i,i+1} \subseteq C_{i,i+1}$ of realizable constraints. We have then to find a vertex ordering for $L_{i+1}$ which minimizes edge crossings while realizing the constraints of $C'_{i,i+1}$. Not all crossing minimization algorithms can be easily adapted to accomplish this task (think, e.g., of the well-known median/barycenter algorithms [4,14]); here we suggest a modification of the penalty minimization technique discussed in [2].

With this method, a penalty digraph $P$ (with vertex set $L_{i+1}$) is built from $B_i$ and $\pi_i$, a feedback arc set $F$ is found on $P$, and a topological sort of the vertices in the subgraph of $P$ obtained by deleting arcs in $F$ is returned as ordering for $L_{i+1}$. We recall that a feedback arc set of a directed graph is a subset of arcs whose removal makes the digraph acyclic; the feedback arc set problem requires to find a minimum weight feedback arc set.

Crossing constraints can be incorporated in the penalty based approach by suitably assigning weights to some arcs of $P$ and by applying the previous algorithm as is. Namely, if $(a, x; b, y)$ is a constraint in $C'_{i,i+1}$ with $a, b \in L_i$ and $x, y \in L_{i+1}$, we add arc $(x, y)$ to $P$ if it does not already exist, and we set...
$w(x, y) = \infty$. Since $C'_{i,i+1}$ is realizable, the subgraph of $P$ induced by the infinite weight arcs is acyclic. Any approximation algorithm for the feedback arc set problem applied on $P$ will therefore choose none of these arcs, thus minimizing crossings while satisfying all the constraints in $C'_{i,i+1}$.

As far as the algorithm above is concerned, it may happen that a pair of edges decomposed by means of dummy vertices may cross several times. This can be avoided by assigning $\infty$ weight to suitably chosen arcs in the constraint digraph.

We also remark that our algorithm can be extended to deal with constraints concerning the relative positions of vertices within each layer. Due to the lack of space, we defer the details to the full paper.

References

On the Validity of Hierarchical Decompositions

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Abstract. Hierarchical decompositions are a useful tool for drawing large graphs. Such decompositions can be represented by means of a data structure called hierarchy tree. In this paper we introduce the notion of $P$-validity of hierarchy trees with respect to a given property $P$: this notion reflects the similarity between the topological structure of the original graph and of any high-level representation of it obtained from the hierarchy tree. We study the $P$-validity when the clustered graph is a tree and property $P$ is the acyclicity, presenting a structure theorem for the $P$-validity of hierarchy trees under these hypotheses.

1 Introduction and Preliminaries

Many graph drawing algorithms have been designed during the last years \cite{2} and most of them move from the following assumption: in order to find a pleasant layout of a graph, we first of all should be able to recognize its graph-theoretical properties, e.g., acyclicity, planarity, bipartiteness, and so on. Actually, exploiting the relevant features of a graph helps produce better visualizations for it. It is also well accepted that only few graph drawing algorithms scale up well and are able to visualize the more and more large graphs that arise in practical applications. Clustering techniques are a useful tool to overcome this drawback and make it possible dealing with graphs not fitting in the screen \cite{3, 4, 5, 6}. Recursively clustering the vertices of a graph leads to a hierarchical decomposition of the graph itself: each cluster in the decomposition is considered as a single node which is visualized instead of a set of vertices of the original graph, considerably reducing the dimension of the drawing. Thanks to the parent relationships between clusters in the hierarchical decomposition, the viewer can move from a high-level representation of the graph to another one by detailing or shrinking clusters. The considerations above immediately lead to the following problem: Given a hierarchical decomposition of a graph $G$ and a graph-theoretical property $P$ satisfied by $G$, does any high-level representation of $G$ obtained from the hierarchical decomposition satisfy property $P$? Though this question naturally arises when using hierarchical decompositions for visualizing large graphs, as far as we

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know it has not been previously studied in literature. Some related work is due to Feng et al., who focus on clustered planar graphs and investigate the concept of compound planarity \[5\].

In this paper we introduce the general notion of \(P\)-\textit{validity} of hierarchical decompositions of graphs with respect to a given property \(P\). This notion reflects the similarity between the topological structure of the original graph and of any high-level representation of it and allows us to compare different hierarchical decompositions associated to the same graph. We then focus on the case where the clustered graph is a tree and property \(P\) is the acyclicity and we present a structural characterization of the \(P\)-\textit{validity} under these hypotheses, i.e., conditions on the structure of the clusters necessary and sufficient to guarantee the \(P\)-\textit{validity} of the decomposition.

We start with preliminary definitions and notation. A hierarchical decomposition of a graph can be represented by means of a data structure called \textit{hierarchy tree} (or sometimes \textit{cluster} or \textit{inclusion tree}) \[1,5\]. A hierarchy tree \(HT = (N, A)\) associated to a graph \(G = (V, E)\) is a rooted tree whose leaves are the vertices of \(G\). Nodes of a hierarchy tree are called \textit{clusters}: a cluster \(c\) represents a set \(V_c\) of vertices of \(G\), namely, the vertices that are the leaves of the subtree rooted at \(c\). We say that such vertices are \textit{covered} by \(c\) and we refer to their number as \textit{cardinality} of \(c\). For brevity, we write \(u \prec c\) to indicate that a vertex \(u \in V\) is covered by a cluster \(c \in N\). We call \textit{singleton cluster} a cluster with cardinality equal to 1. W.l.o.g. we assume that the vertices covered by a cluster \(c\) are a proper subset of the vertices covered by the parent of \(c\) in \(HT\), i.e., in the hierarchy tree there is no node with a unique child. Under this hypothesis, the number of nodes of a hierarchy associated to a \(n\)-vertex graph is at most \(2^n - 1\). This implies that adding a hierarchy tree on the top of a graph requires only space linear in the number of vertices of the graph itself.

It is possible to visit a hierarchy tree in a top-down fashion by performing expand and contract operations on the clusters visualized at any instant of time: these operations define a \textit{navigation} of \(HT\). The previous ideas can be formalized by defining the concepts of \textit{covering} and of \textit{view} \[1,6\]:

\textbf{Definition 1.} A \textit{covering} \(C\) of a graph \(G = (V, E)\) on a hierarchy tree \(HT = (N, A)\) is a subset of the nodes of \(HT\) such that each vertex of \(G\) is covered by exactly a node in \(C\).

A \textit{view} of graph \(G\) on \(HT\) is a graph \(W = (C, L)\) such that \(C\) is a covering of \(G\) on \(HT\) and \(L = \{ (c, c') | c, c' \in C, c \neq c', \exists (u, v) \in E : u \prec c \text{ and } v \prec c' \}\).

We refer to the edges of \(W\) as \textit{links}. It is worth observing that any view is a simple graph: if there exist two edges \((u, v) \in E\) and \((u', v') \in E\) which lead to a connection between two clusters \(c\) and \(c'\), only a single link \((c, c')\) is considered in \(L\). We also call \(W_r\) and \(W_l = G\) the views generated by the root of the hierarchy tree and by all its leaves, respectively; in other words, \(W_r\) and \(W_l\) are the least and the most detailed representations of the graph, respectively.
2 \( \mathcal{P} \)-Validity of Views and Hierarchy Trees

A navigation of a hierarchy tree should help the viewer to focus on particular interesting regions of the graph, according to his/her necessities; hence, if the clusters are generated not taking into account the topological structure of the graph, no benefit may derive for the viewer from the clustering structure. Informally speaking, only vertices in the same “locality” of the graph should be grouped together to form a cluster.

A motivating example of the concept of \( \mathcal{P} \)-validity is illustrated in Figure 1 where two different hierarchical decompositions associated to a chain with 8 vertices are considered (Figure 1(a) and Figure 1(d)). The hierarchy trees \( HT_1 \) (Figure 1(b)) and \( HT_2 \) (Figure 1(e)) corresponding to these decompositions are both complete binary trees of height 3 and differ only in the permutation of their leaves. Figure 1(c) and Figure 1(f) report three views related to coverings \{E, F\}, \{E, C, D\}, and \{A, B, C, D\} in the two decompositions, respectively. All the views built from \( HT_1 \) maintain the structural property of the original graph to be a chain, while the views from \( HT_2 \) lose this property introducing cycles, up to become even a clique.

Fig. 1. Hierarchical decompositions, hierarchy trees, and views of a 8-vertex chain

In order to characterize the “semantic” differences of hierarchy trees with “syntactically” similar or even identical structure it is then natural to introduce a notion of validity of a view and of a hierarchy tree with respect to a certain property \( \mathcal{P} \): from now on we refer to this concept as \( \mathcal{P} \)-validity.

Definition 2. Let \( HT \) be a hierarchy tree associated to a graph \( G \) and let \( \mathcal{P} \) be a property satisfied by \( G \). A view \( W \) of \( G \) obtained from \( HT \) is \( \mathcal{P} \)-valid iff \( W \)
satisfies property \( \mathcal{P} \). \( HT \) is \( \mathcal{P} \)-valid iff all the views of \( G \) obtained from it are \( \mathcal{P} \)-valid.

In view of the fact that Definition 2 is parametric in the property \( \mathcal{P} \) to be considered, different notions of validity may be thought for different classes of graphs. For example, we may require a view of a bipartite graph to be bipartite or a view of a planar graph to be planar. But even more sophisticated definitions for \( \mathcal{P} \) can be considered: as an example, we refer to the \( c \)-planarity property introduced by Feng et al. [5]. The interest in characterizing \( \mathcal{P} \)-valid hierarchy trees naturally follows from the considerations in Section 1. In this paper we focus on the case where the clustered graph is a tree and property \( \mathcal{P} \) is the acyclicity: for brevity, we will speak of valid views/hierarchy trees under these hypotheses. In Section 3 we prove necessary and sufficient conditions for a hierarchy tree to be valid.

3 A Structural Characterization of Valid Hierarchy Trees

Before presenting the structure theorem for the validity of a hierarchy tree, we give preliminary definitions and lemmas useful for proving it. First of all we study the connectivity of views obtained from hierarchy trees of connected graphs.

Lemma 1. Each view obtained from a hierarchy tree associated to a connected graph is connected.

In the following we denote with \( S(c) \) the subgraph of \( T = (V, E) \) induced by the vertices covered by a node \( c \) of \( HT \). For each \( u, v \in T \), let \( \text{path}_T(u,v) \) be the path joining \( u \) and \( v \) in \( T \) and let \( \text{dist}_T(u,v) \) be the length of this path.

Definition 3. Let \( c \) be a node of a hierarchy tree \( HT \) associated to a free tree \( T \). Let \( u \) and \( v \) be two vertices of \( T \) covered by \( c \). \( u, v \) are a broken pair of cluster \( c \) iff they are neither coincident nor connected in \( S(c) \). A broken pair \( u, v \) is a minimum-distance broken pair of \( c \) iff \( w \not\prec c \), \( \forall w \in \text{path}_T(u,v) \) such that \( w \neq u, v \).

Lemma 2. Let \( W \) be a view on a hierarchy tree \( HT \) associated to a free tree \( T \). If \( W \) is not acyclic, then in each cycle \( C \) there is at least a cluster containing a broken pair.

Proof. Let \( C = (c_1, \ldots, c_h) \) be a cycle in \( W \). Each cluster \( c_i \in C \) is endpoint of two links \( (c_{i-1}, c_i) \) and \( (c_i, c_{i+1}) \), respectively (to simplify the notation we assume that \( h + 1 = 1 \)). Let us call \( (r_{i-1}, l_i) \) and \( (r_{i}, l_{i+1}) \) two tree edges which derive links \( (c_{i-1}, c_i) \) and \( (c_i, c_{i+1}) \), respectively. For each cluster \( c_i \) we therefore identify two vertices covered by it, named \( l_i \) and \( r_i \) (see Figure 2). If no broken pair exists in any cluster of \( C \), for each \( i \in [1, h] \) vertices \( l_i \) and \( r_i \) are either coincident or connected in \( S(c_i) \). This implies the existence of a cycle in tree \( T \), that is a contradiction.
Theorem 1. Let \( T = (V, E) \) be a free tree and let \( HT = (N, A) \) be a hierarchy tree associated to \( T \). \( HT \) is valid if and only if for each minimum-distance broken pair \( u, v \) of \( HT \), \( \text{dist}_T(u, v) = 2 \).

Proof. We first prove the necessary condition, showing that the existence in \( HT \) of a minimum-distance broken pair \( u, v \) such that \( \text{dist}_T(u, v) > 2 \) implies that \( HT \) is not valid. Let \( c \) be a node of \( HT \) such that \( u, v \) are a minimum-distance broken pair in \( c \). Let us consider the view \( W = (C, L) \) where covering \( C \) consists only of singleton clusters, except for cluster \( c: C = \{c\} \cup \{\{x\}, x \in V \text{ and } x \not\prec c\} \). Let \( z \) and \( w \) be the vertices in path \( T(u, v) \) such that \( (u, z) \in E \) and \( (w, v) \in E \). Since \( u, v \) is a minimum-distance broken pair of \( c \), \( z \not\prec c \), \( w \not\prec c \), and any other vertex in path \( T(u, v) \) is not covered by \( c \). Moreover, since \( \text{dist}_T(u, v) > 2 \), \( \text{dist}_T(z, w) > 0 \), i.e., \( z \) and \( w \) are distinct vertices. Hence \( \langle c, \{z\}, \ldots, \{w\}, c \rangle \) is a cycle in \( W \), proving that \( HT \) is not valid.

We now focus on the sufficient condition, proving that a contradiction can be derived if we suppose that \( HT \) is not valid while satisfying the property on the minimum-distance broken pairs formulated in the statement of the theorem. Namely, we assume that in \( HT \) there exists a non-valid view \( W = (C, L) \), i.e., a view \( W \) that is not a tree. Since \( W \) must be connected due to Lemma 1, it must contain a simple cycle \( C \). We prove that this assumption leads to a contradiction.

The general idea of the proof is to convert \( W \) into another view \( W' \), still existing on the hierarchy tree, such that: (a) the number of singletons of \( W' \) is strictly greater than the number of singletons of \( W \); (b) the cycle \( C \) of \( W \) is also changed into a simple cycle \( C' \) in \( W' \). The sequence of manipulations that we perform has finite length, as the number of singletons is clearly upper bounded by \( n = |V| \). We can therefore prove that at some step during this process we find a contradiction due to Lemma 2, since we obtain a cycle with no broken pairs.

Let \( C = (c_1, \ldots, c_h) \) be a simple cycle of length \( h \) in \( W \) and let \( c_i, l_i, \) and \( r_i \) be defined as in the proof of Lemma 2 for \( 1 \leq i \leq h \) (see also Figure 2). Due to Lemma 2, a broken pair must exist in \( C \). Let \( l_i \) and \( r_i \) be the vertices in such a broken pair of \( C \). We remark that \( l_i \) and \( r_i \) are not necessarily a minimum-distance broken pair.

Let us consider the path from \( l_i \) to \( r_i \) in \( T \), which is unique and not completely contained in \( S(c_i) \). On this path we can univocally identify a set of \( k \) vertices \( u_j \), for \( 1 \leq j \leq k \), such that \( u_j \prec c_i \) but its successor on path \( T(l_i, r_i) \) is not covered by \( c_i \). Analogously, we can univocally identify a set of \( k \) vertices \( v_j \), for \( 1 \leq j \leq k \), such that \( v_j \prec c_i \) but its predecessor on path \( T(l_i, r_i) \) is not covered.
by \(c_i\). Observe that it could be \(u_j = v_j\) for some \(j\), but vertices with different indexes always belong to different connected components of \(S(c_i)\).

It is clear that \(v_1 = l_i\) and \(u_k = r_i\). Besides, it is easy to see that the pairs \(u_j, v_{j+1}\), for \(1 \leq j \leq k - 1\), are minimum distance broken pairs of cluster \(c_i\) and, by hypothesis, we know that \(\text{dist}_T(u_j, v_{j+1}) = 2\). Let us call \(z_j\) the unique vertex of \(T\) in the path between \(u_j\) and \(v_{j+1}\) and let \(c_j'\) be the cluster of \(W\) that covers \(z_j\). The configuration is illustrated in Figure 3(a).

We now change view \(W\) into \(W'\) by expanding cluster \(c_i\) at the singleton level, i.e., by substituting \(c_i\) with the set of singleton clusters corresponding to the vertices in \(S(c_i)\). In the following we prove that we are able to exhibit in \(W'\) a simple cycle \(C'\). Let us first consider the cycle \(\hat{C}\) shown in Figure 3(b). \(\hat{C}\) clearly exists in \(W'\), being obtained from \(\mathcal{C}\) by unrolling \(\text{path}_T(l_i, r_i)\) (compare vertices and clusters involved in Figure 3(a) and in Figure 3(b), respectively). However, \(\hat{C}\) is not necessarily simple. W.l.o.g. we can assume that \(c_j' \neq c_s' \forall j, s \in [1, k - 1], j \neq s\). We can always reduce to this situation as follows: while \(\hat{C}\) contains a pair \(j, s\) such that \(1 \leq j < s \leq k - 1\) and \(c_j' = c_s'\), substitute the path \(\langle c_j', \ldots, c_s' \rangle\) with the path \(\langle c_j' \rangle\). After this operation all the \(c_j'\) are distinct. If \(\forall j \in [1, k - 1], c_j' \notin \mathcal{C}, C' = \hat{C}\) is a simple cycle. Otherwise \(\exists j \in [1, k - 1]\) such that \(c_j' \in \mathcal{C}\) and two cases may happen:

- \(c_1' = c_{i-1}'\): change \(\hat{C}\) by replacing the non-simple path \(\langle c_{i-1}, p_1, c_1', p_2 \rangle\) with the simple path \(\langle c_{i-1}, p_2 \rangle\). If \(k = 1\) or \(\forall j \in [2, k - 1], c_j' \notin \mathcal{C}\), then the modified \(\hat{C}\) is a simple cycle and we can choose \(C' = \hat{C}\). Otherwise, let \(s\) be the smallest value in \([2, k - 1]\) such that \(c_s' \in \mathcal{C}\), i.e., \(c_s' = c_t\) for some \(t \in [1, h]\), \(t \neq i\). Moving clockwise on \(\hat{C}\), we find the simple cycle \(C' = \langle c_{i-1}, p_2, c_2', \ldots, p_s, c_s' = c_t \sim c_{i-1} \rangle\), where \(\sim\) indicates a subpath of \(p\).

- \(c_1' \neq c_{i-1}'\): let \(s\) be the smallest value in \([1, k - 1]\) such that \(c_s' \in \mathcal{C}\), i.e., \(c_s' = c_t\) for some \(t \in [1, h]\), \(t \neq i\). If \(s = 1\) then \(C' = \langle c_{i-1}, p_1, c_1', c_2', \ldots, p_s, c_s' = c_t \sim c_{i-1} \rangle\) is a simple cycle in \(W'\). Note that the length of the path \(\langle c_t \sim c_{i-1} \rangle\) is \(\geq 1\), because \(c_t = c_1' \neq c_{i-1}'\). If \(s > 1\) then let \(C' = \langle c_{i-1}, p_1, c_1', p_2, c_2', \ldots, p_s, c_s' = c_t \sim c_{i-1} \rangle\).
In this case it could be \( c_i = c_{i-1} \); however, the fact that the path from \( c_{i-1} \) to \( c'_s \) has length \( \geq 3 \) guarantees \( \mathcal{C}' \) to be a cycle of \( W' \). By construction we also know that \( \mathcal{C}' \) is simple.

In any case we are therefore able to find a view \( W' \) containing more singletons than \( W \) and to exhibit a simple cycle \( \mathcal{C}' \) in \( W' \). Iterating this reasoning, we obtain either a cycle with no broken pairs, which is absurd due to Lemma 2 or a cyclic view containing only singleton clusters, i.e., \( W_i \). This is also a contradiction because \( W_i \) is acyclic being equal to \( T \).

Theorem 1 implies that the validity of a hierarchy tree can be checked in polynomial time. We remark that this does not immediately follow from Definition 2 since the number of views in a hierarchy tree may be exponential.

References

Lower Bounds on the Minus Domination and \( k \)-Subdomination Numbers

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Abstract. A three-valued function \( f \) defined on the vertex set of a graph \( G = (V, E) \), \( f : V \to \{-1, 0, 1\} \) is a minus dominating function if the sum of its function values over any closed neighborhood is at least one. That is, for every \( v \in V \), \( f(N[v]) \geq 1 \), where \( N[v] \) consists of \( v \) and all vertices adjacent to \( v \). The weight of a minus function is \( f(V) = \sum_{v \in V} f(v) \).

The minus domination number of a graph \( G \), denoted by \( \gamma^{-}(G) \), equals the minimum weight of a minus dominating function of \( G \). In this paper, sharp lower bounds on minus domination of a bipartite graph are given. Thus, we prove a conjecture proposed by J. Dunbar et al. (Discrete Math. 199(1999) 35-47), and we give a lower bound on \( \gamma_{ks}(G) \) of a graph \( G \).

1 Introduction

For a graph \( G = (V, E) \) with vertex set \( V \) and edge set \( E \), the open neighborhood of \( v \in V \) is \( N(v) = \{ u \in V | uv \in E \} \) and the closed neighborhood of \( v \) is \( N[v] = \{ v \} \cup N(v) \). For a set \( S \) of vertices, we define the open neighborhood \( N(S) = \bigcup_{v \in S} N(v) \), and the closed neighborhood \( N[S] = N(S) \cup S \). A dominating set \( S \) for a graph \( G = (V, E) \) is a subset of the vertex set \( V \) such that every vertex \( v \in V \) is either in \( S \) or adjacent to a vertex in \( S \). The domination number of \( G \), \( \gamma(G) \), equals the minimum cardinality of a dominating set.

For a real function \( f \) defined on vertices of a graph \( G \) and \( S \subseteq V \), write \( f(S) = \sum_{v \in S} f(v) \) and \( f[v] = f(N[v]) \). A minus dominating function of \( G \) is defined in [3] as a function \( f : V \to \{-1, 0, 1\} \) such that \( f[v] \geq 1 \) for each \( v \in V \). A signed dominating function of \( G \) is defined in [5] as \( f : V \to \{-1, 1\} \) satisfying \( f[v] \geq 1 \) for all \( v \in V \). A minus(signed) dominating function \( f \) is minimal if every minus(signed) dominating function \( g \) satisfying \( g(v) \leq f(v) \) for every \( v \in V \), is equal to \( f \). It is easy to see that a minus dominating function is minimal if and only if for every vertex \( v \in V \) with \( f[v] \geq 0 \), there exists a vertex \( u \in N[v] \) with \( f[u] = 1 \) and a signed function is minimal if and only if every vertex \( v \) of weight 1, there exists some \( u \in N[v] \)
such that \( f[u] = 1 \) or \( 2 \). The \textit{minus domination number} for a graph \( G \) is 
\[
\gamma^-(G) = \min \{ f(V) \mid f \text{ is a minimal minus dominating function} \}.
\]
Likewise the \textit{signed domination number} for a graph \( G \) is 
\[
\gamma_s(G) = \min \{ f(V) \mid f \text{ is a minimal signed dominating function} \}
\]

A \textit{majority dominating function} of \( G \) is define in [2] as \( f : V \rightarrow \{-1, 1\} \) such that \( f[v] \geq 1 \) for at least half the vertices of \( G \), and the minimum weight of such a function is the \textit{majority domination number}.

For a positive integer \( k \), a \textit{k-subdominating function}(kSF) of \( G \) is a function \( f : V \rightarrow \{-1, 1\} \) such that \( f[v] = \sum_{u \in N(v)} f(u) \geq 1 \) for at least \( k \) vertices of \( G \). The aggregate \( ag(f) \) of such a function is defined by \( ag(f) = \sum_{v \in V} f(v) \) and the \textit{k-subdomination number} \( \gamma_{ks}(G) \) by \( \gamma_{ks} = \min \{ ag(f) : f \text{ is a kSF of } G \} \). In the special cases \( k = |V| \) and \( k = \lceil \frac{|V|}{2} \rceil \), \( \gamma_{ks} \) is respectively the signed domination number \( \gamma_s(G) \) and the majority domination number \( \gamma_{maj}(G) \).

Since the problems of determining the signed domination number and minus domination number are NP-complete, many works on bounds for \( \gamma^-(G) \) and \( \gamma_s(G) \) were studied in [4, 7, 8, 9, 10, 11, 12]. In [3], the following conjecture was given.

\textbf{Conjecture 1} [3]. If \( G \) is a bipartite graph of order \( n \), then \( \gamma^-(G) \geq 4(\sqrt{n+1}-1) - n \)

\section{Lower Bounds on Minus Domination of a Bipartite Graph}

\textbf{Theorem 1}. If \( G = (X, Y) \) is a bipartite graph, then 
\[
\gamma^-(G) \geq 4(\sqrt{n+1}-1) - n
\]

\textbf{Proof}. Let \( f \) be a minus dominating function of \( G \) satisfying \( f(V) = \gamma^-(G) \) and

\[
M = \{ v \in V \mid f(v) = -1 \}
\]
\[
P = \{ v \in V \mid f(v) = 1 \}
\]
\[
Q = \{ v \in V \mid f(v) = 0 \}
\]

\[
M_X = M \cap X, \quad M_Y = M \cap Y, \quad P_X = P \cap X, \quad P_Y = P \cap Y, \quad Q_X = Q \cap X, \quad Q_Y = Q \cap Y, \quad m_x = |M_X|, \quad m_y = |M_Y|, \quad p_x = |P_X|, \quad p_y = |P_Y|, \quad q_x = |Q_X|, \quad q_y = |Q_Y|.
\]

Since \( f[v] \geq 1 \) for every \( v \in V \), we have \( |N(v) \cap P_X| \geq 2 \) for every \( v \in M_Y \). So

\[
e(P_X, M_Y) \geq 2m_y.
\]

For every \( v \in P_X \), \( |N(v) \cap M_Y| \leq |N(v) \cap P_Y| \). Then

\[
e(P_X, M_Y) = \sum_{v \in P_X} |N(v) \cap M_Y|
\]
\[
\leq \sum_{v \in P_X} |N(v) \cap P_Y|
\]
\[
\leq p_x p_y.
\]
By (1) and (2) we have

\[ 2m_y \leq pxpy. \]

Similarly

\[ 2m_x \leq pxpy, \]

then

\[ m_x + m_y \leq pxpy. \] (3)

Since

\[ n = qx + qy + m_x + m_y + px + py, \]

and

\[ 2\sqrt{pxpy} \leq px + py, \]

we have

\[ 2\sqrt{pxpy} + m_x + m_y + qx + qy \leq n. \] (4)

Using (3) and (4) we have

\[ 2\sqrt{m_x + m_y} + m_x + m_y + qx + qy \leq n, \] (5)

and

\[ 2\sqrt{m_x + m_y} + m_x + m_y \leq n. \] (6)

Using (5) we obtain

\[ \gamma^{-}(G) = f(V(G)) \]

\[ = n - (qx + qy) - 2(m_x + m_y) \]

\[ \geq m_x + m_y + 2\sqrt{m_x + m_y} - 2(m_x + m_y) \]

\[ = 2\sqrt{m_x + m_y} - (m_x + m_y). \] (7)

For notation convenience, we define the following

\[ a = \sqrt{m_x + m_y}, \]

\[ h(y) = y^2 + 2y \quad (y \geq 1), \]

\[ g(y) = 2y - y^2 \quad (y \geq 1). \]

Since \( \frac{dh}{dy} = 2y + 2 \geq 2, \frac{dg}{dy} = 2 - 2y \leq 0 \), so \( h(y) \) is a monotonous increasing function and \( g(y) \) is a monotonous decreasing function. By (6) we have \( h(a) = a^2 + 2a \leq n \). And when \( y = -1 + \sqrt{1 + n} \),

\[ h(y) = (-1 + \sqrt{1 + n})^2 + 2(-1 + \sqrt{1 + n}) \]

\[ = 1 - 2\sqrt{1 + n} + 1 + n - 2 + 2\sqrt{1 + n} \]

\[ = n. \]
So \( a \leq -1 + \sqrt{1+n} \).

By (7) we obtain

\[
\gamma^-(G) \geq g(a)
\geq g(-1 + \sqrt{1+n}) \\
= 2(-1 + \sqrt{1+n}) - (-1 + \sqrt{1+n})^2 \\
= 2(-1 + \sqrt{1+n}) - (1 - 2\sqrt{1+n} + 1 + n) \\
= 4(\sqrt{n+1} - 1) - n.
\]

We now show that this bound is best possible by the following graphs \( G \) construct by J. Dunbar et al.\[3\]. Let \( s \geq 4 \) be an even integer, and let \( H \) be isomorphic to \( \frac{s}{2} \) disjoint copies of \( K_{2,s} \). Let \( H_1 \) and \( H_2 \) be two disjoint copies of \( H \). Further, let \( X_i \) and \( Y_i \) be the sets of vertices of \( H_i \) of degree 2 and \( s \), respectively, for \( i = 1, 2 \). Now let \( G \) be the graph obtained from \( H_1 \cup H_2 \) by joining every vertex of \( Y_1 \) to every vertex of \( Y_2 \). Then \( G \) is a bipartite graph of order \( n = s(s+2) \) with partite sets \( X_1 \cup Y_2 \) and \( X_2 \cup Y_1 \). Let \( f \) be the function on \( G \) defined as follows: let \( f(v) = -1 \) if \( v \in X_1 \cup X_2 \), and let \( f(v) = 1 \) if \( v \in Y_1 \cup Y_2 \). Then it is easy to verify that \( f \) is a minus dominating function on \( G \) with \( \gamma^-(G) = f(V(G)) = 2s - s^2 = 4(\sqrt{n+1} - 1) - n \). \(\square\)

**Theorem 2.** If \( G = (X,Y) \) is a bipartite graph of order \( n \), then

\[
\gamma^-(G) \geq \left[ n - \left( \frac{\epsilon}{\delta} + \frac{\epsilon}{1+\max(\delta_X,\delta_Y)} \right) \right]
\]

where \( \delta_X = \min\{d(v)|v \in X\}, \delta_Y = \min\{d(v)|v \in Y\} \), and the bound is sharp.

**Proof.** Let \( f, M_X, M_Y, P_X, P_Y, Q_X, Q_Y, m_x, m_y, q_x, q_y, p_x \) and \( p_y \) be defined as in the proof of Theorem 1. For any \( x \in V \), let \( t_x \) denotes the number of vertices of weight 0 in \( N(x) \). Then we have

\[
|N(x) \cap M| \leq \begin{cases} 
\frac{d(x)-t_x}{2}, & \text{if } x \in P \\
\frac{d(x)-1-t_x}{2}, & \text{if } x \in Q \\
\frac{d(x)-t_x}{2} - 1 & \text{if } x \in M
\end{cases}
\]

So

\[
\sum_{y \in M_Y} d(y) = \sum_{x \in P_X} |N(x) \cap M_Y| + \sum_{x \in Q_X} |N(x) \cap M_Y| + \sum_{x \in M_X} |N(x) \cap M_Y|
\leq \sum_{x \in P_X} \frac{d(x)-t_x}{2} + \sum_{x \in Q_X} \frac{d(x)-1-t_x}{2} + \sum_{x \in M_X} \left( \frac{d(x)-t_x}{2} - 1 \right)
\]

\[
= \sum_{x \in X} \left( \frac{d(x)}{2} - \frac{t_x}{2} \right) - \frac{1}{2} q_x - m_x
\]

(8)

Obviously

\[
m_y \delta_Y \leq \sum_{y \in M_Y} d(y)
\]

(9)
\[
\sum_{x \in X} \frac{t_x}{2} = \frac{1}{2} \sum_{y \in Q_Y} d(y) \geq \frac{1}{2} \delta_Y q_y. \tag{10}
\]

Combining (8), (9) and (10) we obtain
\[
(q_y + 2m_y)\delta_Y + (q_x + 2m_x) \leq \epsilon. \tag{11}
\]
Similarly we have
\[
(q_x + 2m_x)\delta_X + (q_y + 2m_y) \leq \epsilon. \tag{12}
\]
If \(q_x + 2m_x \leq q_y + 2m_y\), by (11) and (12) we have
\[
q_x + 2m_x \leq \frac{\epsilon}{1 + \max(\delta_X, \delta_Y)},
\]
\[
q_y + 2m_y \leq \frac{\epsilon}{\delta_Y}.
\]
So
\[
\gamma^-(G) = n - (q_x + 2m_x + q_y + 2m_y)
\geq \left\lceil n - \left(\frac{\epsilon}{\delta_Y} + \frac{\epsilon}{1 + \max(\delta_X, \delta_Y)}\right) \right\rceil
\geq \left\lceil n - \left(\frac{\epsilon}{\delta} + \frac{\epsilon}{1 + \max(\delta_X, \delta_Y)}\right) \right\rceil.
\]
If \(q_y + 2m_y < q_x + 2m_x\), by (11) and (12) we have
\[
q_y + 2m_y \leq \frac{\epsilon}{1 + \max(\delta_X, \delta_Y)},
\]
\[
q_x + 2m_x \leq \frac{\epsilon}{\delta_X}.
\]
So
\[
\gamma^-(G) = n - (q_x + 2m_x + q_y + 2m_y)
\geq \left\lceil n - \left(\frac{\epsilon}{\delta_X} + \frac{\epsilon}{1 + \max(\delta_X, \delta_Y)}\right) \right\rceil
\geq \left\lceil n - \left(\frac{\epsilon}{\delta} + \frac{\epsilon}{1 + \max(\delta_X, \delta_Y)}\right) \right\rceil.
\]
In fact, this bound is sharp, it is easy to check that \(\gamma^-(K_{1,k}) = 1 = \left\lceil n - \frac{\epsilon}{\delta} + \frac{\epsilon}{1 + \max(\delta_X, \delta_Y)} \right\rceil\).

3 A Lower Bound on \(k\)-Subdomination Number of a Graph

The concept of \(k\)-subdomination was introduced by Cockayne et al.[2]. In [2], Cockayne et al. established a sharp lower bound on \(\gamma_{ks}\) for trees. Moreover,
they also gave a sharp lower bound on $\gamma_{ks}$ for trees if $k \leq \frac{n}{2}$ and proposed a conjecture.

**Theorem 3 [2].** For any $n$-vertex tree $T$ and integer $k \in \{1, 2, \ldots, n\}$, $\gamma_{ks} \leq 2(k + 1) - n$.

**Conjecture 2 [2].** For any $n$-vertex tree and any $k$ with $\frac{1}{2}n < k \leq n$, $\gamma_{ks} \leq 2k - n$.

In [13], the conjecture was proved and an upper bound for a connected graph was given.

**Theorem 4 [13].** For any connected graph of order $n$ and any $k$ with $\frac{1}{2}n < k \leq n$, then

$$\gamma_{ks} \leq 2 \left\lceil \frac{k}{n - k + 1} \right\rceil (n - k + 1) - n.$$ 

In this section, we give a lower bound for a graph $G$.

**Theorem 5.** For any graph $G$ of order $n$ and size $\epsilon$,

$$\gamma_{ks} \geq n - \frac{2\epsilon + (n - k)(\Delta + 2)}{\delta + 1}.$$

**Proof.** Let $f$ be a $k$-subdominating function on $G$ with $f(V) = \gamma_{ks}(G)$. Let $P$ and $M$ be the sets of vertices in $G$ that are assigned the values 1 and $-1$, respectively. Then $|P| + |M| = n$ and $\gamma_{ks}(G) = |P| - |M| = n - 2|M|$. Furthermore, we let

$$P_1 = \{v \in P | f[v] \geq 1\}$$
$$P_2 = P - P_1$$
$$M_1 = \{v \in M | f[v] \geq 1\}$$
$$M_2 = M - M_1$$

Clearly, $|P_1| + |M_1| \geq k$. Since each vertex $v$ of $P_1$ is adjacent to at most $\frac{1}{2}d(v)$ vertices of $M$, each vertex $v$ of $M_1$ is adjacent to at most $\frac{1}{2}d(v) - 1$ vertices of $M$. We have

$$\delta|M| \leq \sum_{v \in M} d(v) = \sum_{v \in V} |M \cap N(v)|$$
$$\leq \sum_{v \in P_1} \frac{d(v)}{2} + \sum_{v \in M_1} \left(\frac{d(v)}{2} - 1\right) + \sum_{v \in P_2 \cup M_2} d(v)$$
$$= \frac{1}{2} \sum_{v \in V} d(v) - |M_1| + \frac{1}{2} \sum_{v \in P_2 \cup M_2} d(v)$$
$$\leq \epsilon - |M| - \frac{1}{2} \sum_{v \in P_2 \cup M_2} (d(v) + 2)$$
$$\leq \epsilon - |M| + (|P_2| + |M_2|) \frac{\Delta + 2}{2}.$$
As $|P_2| + |M_2| \leq n - k$, it follows that

$$|M| \leq \frac{2\epsilon + (n - k)(\Delta + 2)}{2(\delta + 1)}$$

Thus

$$\gamma_{ks} = n - 2|M|$$

$$\geq n - \frac{2\epsilon + (n - k)(\Delta + 2)}{\delta + 1}.$$ 

This completes the proof of Theorem 5.

For the graphs in which each vertex has odd degree, the lower bound on $\gamma_{ks}$ in Theorem 5 can be improved slightly.

**Theorem 6.** For every graph $G$ in which each vertex has odd degree,

$$\gamma_{ks} \geq n - \frac{2\epsilon + (n - k)(\Delta + 2) - k}{\delta + 1}.$$ 

**Proof.** Let $f, P, M, P_1, P_2, M_1$ and $M_2$ be defined as in the proof of Theorem 5.

Since every vertex of $G$ has odd degree, it is easy to see that each vertex $v$ of $P_1$ is adjacent to at most $\frac{d(v) - 1}{2}$ vertices of $M$, each vertex $v$ of $M_1$ is adjacent to at most $\frac{d(v) - 1}{2} - 1$ vertices of $M$. Hence we have

$$\delta|M| \leq \sum_{v \in M} d(v) = \sum_{v \in V} |M \cap N(v)|$$

$$\leq \sum_{v \in P_1} \frac{d(v) - 1}{2} + \sum_{v \in M_1} \left( \frac{d(v) - 1}{2} - 1 \right) + \sum_{v \in P_2 \cup M_2} d(v)$$

$$= \frac{1}{2} \sum_{v \in V} d(v) - \frac{1}{2}(|P_1| + |M_1|) - |M_1| + \frac{1}{2} \sum_{v \in P_2 \cup M_2} d(v)$$

$$\leq \epsilon - \frac{1}{2}(|P_1| + |M_1|) - |M| + \frac{1}{2} \sum_{P_2 \cup M_2} d(v) + \frac{\Delta + 2}{2}.$$ 

Since $|P_1| + |M_1| \geq k$, $|P_2| + |M_2| \leq n - k$, we have

$$(\delta + 1)|M| \leq \epsilon - \frac{k}{2} + \frac{(\Delta + 2)(n - k)}{2}.$$ 

Hence

$$|M| \leq \frac{2\epsilon + (\Delta + 2)(n - k) - k}{2(\delta + 1)}.$$ 

Thus

$$\gamma_{ks}(G) = n - 2|M|$$

$$\geq n - \frac{2\epsilon + (\Delta + 2)(n - k) - k}{\delta + 1}.$$
This completes the proof of Theorem 6. □

By Theorem 5 and Theorem 6, we easily obtain the following lower bounds on $\gamma_{ks}$ for $r$-regular graphs.

**Theorem 7.** Let $G$ be a $r$-regular graph of order $n$, then

$$\gamma_{ks} \geq \begin{cases} \frac{r+2}{r+1}k - n, & \text{for } r \text{ even} \\ \frac{r+3}{r+1}k - n, & \text{for } r \text{ odd} \end{cases}$$

In the special cases where $k = |V|$ and $k = \lceil \frac{|V|}{2} \rceil$, Theorem 7 deduces to the following results.

**Corollary 7** [9]. For every $r$-regular graph $G$ of order $n$,

$$\gamma_s(G) \geq \begin{cases} \frac{2n}{r+3}, & \text{for } r \text{ odd} \\ \frac{n}{r+1}, & \text{for } r \text{ even} \end{cases}$$

and the bounds are sharp.

**Corollary 8** [9]. For every $r$-regular $(r \geq 2)$ graph $G$ of order $n$,

$$\gamma_{maj}(G) \geq \begin{cases} \frac{1-r}{2(r+1)}n, & \text{for } r \text{ odd} \\ \frac{-r}{2(r+1)}n, & \text{for } r \text{ even} \end{cases}$$

and the bounds are sharp.

**References**


Edge Connectivity vs Vertex Connectivity in Chordal Graphs

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Abstract. It is well known that in a graph, \( \kappa(G) \leq \lambda(G) \leq \delta(G) \), where \( \kappa(G) \), \( \lambda(G) \) and \( \delta(G) \) denote the vertex connectivity, edge connectivity and the minimum degree of \( G \), respectively. We show that in chordal graphs, if \( \lambda(G) \neq \delta(G) \), then \( \lambda(G) \geq 2\kappa(G) - 1 \). In contrast, in a general graph \( \lambda(G) \) can be equal to \( \kappa(G) \).

1 Introduction

1.1 Chordal Graphs: Some Definitions and Notation

Let \( C \) be a cycle in a graph \( G \). A chord of \( C \) is an edge of \( G \) joining two vertices of \( C \) which are not consecutive. A graph \( G \) is called a chordal (or triangulated) graph iff every cycle in \( G \) of length 4 or more has a chord. Chordal graphs arise in many applications (see [3,4,5]). Chordal graphs constitute one of the most important subclasses of perfect graphs [3].

Let \( G = (V, E) \) be a connected undirected graph. Throughout this paper we will use \( V \) for the set of vertices of \( G \) and \( E \) for the set of edges. \( |V| \) will be denoted by \( n \). \( N(v) \) will denote the set of neighbours of \( v \), that is \( N(v) = \{u \in V : (u, v) \in E\} \). For \( A \subseteq V \), we use \( N(A) \) to denote the set \( \bigcup_{v \in A} N(v) - A \). The subgraph of \( G \) induced by the nodes in \( A \) will be denoted by \( G(A) \).

A bijection \( f : V \to \{1, 2, \cdots, n\} \) is called an ordering of the vertices of \( G \). Then \( f(v) \) is referred to as the number associated with the vertex \( v \), or simply the number of \( v \) with respect to the ordering \( f \). Given an ordering \( f \) of a graph \( G \), we define the following terms.

Definition 1. Let \( A \subseteq V \). The highest \( (A) \) is defined to be the vertex with the highest number in \( A \).

Definition 2. A path \( P = (w_1, w_2, \cdots, w_k) \) in \( G \) is called an increasing path, iff \( f(w_1) < f(w_2) < \cdots < f(w_k) \). It is called a decreasing path iff \( f(w_1) > f(w_2) > \cdots > f(w_k) \). A single node can be considered as either increasing or decreasing.
Definition 3. A vertex $u \in N(v)$ is called a higher neighbour of $v$ iff $f(u) > f(v)$. The set of higher neighbours of $v$ will be denoted by $N_h(v)$, i.e.

$$N_h(v) = \{u \in N(v) : f(u) > f(v)\}$$

Definition 4. An ordering $f$ of $G$ is called a perfect elimination ordering (PEO) iff for each $v \in V$, $G(\{v\} \cup N_h(v))$ is a complete subgraph (clique) of $G$. Then $f(v)$ will be called the PEO number of $v$.

Note that every graph may not have a PEO. The following theorem is well known (see [3]).

Theorem 1. An undirected graph $G$ is chordal if and only if there exists a PEO for $G$.

Note that there can be more than one PEO for a given chordal graph. We just use $PEO(v)$ to denote the PEO number of $v$, since our arguments are valid with respect to any PEO. Whenever we talk of a chordal graph, we will assume a PEO for it and the words highest, increasing path etc will be used with the respect to this PEO.

1.2 Edge Connectivity and Vertex Connectivity

The vertex connectivity $\kappa(G)$ of an undirected graph is defined to be the minimum number of vertices whose removal results in a disconnected graph or a trivial graph (i.e., single node). A subset $S \subset V$ is called a separator if $G(V - S)$ has at least two components. Then, $\kappa(G)$ is the size of the minimum sized separator. Note that a complete graph has no separator and its (vertex) connectivity is by definition $n - 1$.

The following result called Menger’s theorem is well known among graph theorists. An elegant proof due to Dirac is given in [2].

Theorem 2. Menger’s theorem: The minimum number of vertices separating two nonadjacent nodes $s$ and $t$ is equal to the maximum number of vertex disjoint $s - t$ paths.

Thus, there are at least $\kappa(G)$ vertex disjoint paths between every pair of nonadjacent nodes of $G$.

The edge connectivity $\lambda(G)$ of a graph $G$ is defined to be the minimum number of edges whose removal results in a disconnected graph or a trivial graph. A minimum cut is a cut with its size $= \lambda(G)$. Note that the minimum degree $\delta(G)$ of the graph is an upper bound for mincut size (i.e., edge connectivity).

1.3 Our Result

It is well known that the following inequality (due to Whitney [6]) holds in a general undirected graph.

$$\kappa(G) \leq \lambda(G) \leq \delta(G)$$
where $\delta(G)$ is the minimum degree of the graph. For a discussion of this inequality and for related work see the chapter on connectivity in [2]. In [1], Chartrand and Harary show that for all integers $a, b, c$, such that $0 < a \leq b \leq c$, there exists a graph $G$ with $\kappa(G) = a$, $\lambda(G) = b$ and $\delta(G) = c$.

In this paper, we study this inequality when restricted to the class of chordal graphs. We show that there is a “gap” between $\lambda(G)$ and $\kappa(G)$ in chordal graphs provided $\lambda(G) \neq \delta(G)$. More specifically, if $\lambda(G) \neq \delta(G)$, then $\lambda(G) \geq 2\kappa(G) - 1$.

2 Edge Connectivity vs Vertex Connectivity in Chordal Graphs

First we make some simple observations about mincuts in undirected graphs. Remember that mincut is a minimum collection of edges which when removed disconnects the graph. Let $V$ be partitioned to nonempty disjoint sets $A$ and $B$ by a cut. We denote the set of edges in the cut by $(A, B)$.

Lemma 1. Let $(A, B)$ be a mincut of a connected undirected graph $G$. Then $G(A)$ and $G(B)$ are connected.

**Proof :** Suppose that $G(A)$ is not connected. Let $G(A_1)$ be a connected component of $G(A)$, where $A_1 \subset A$. Let $A_2 = A - A_1$. Clearly $(A_1, B \cup A_2)$ is a cut of $G$ and $|(A_1, B \cup A_2)| < |(A, B)|$, since $(A_1, B \cup A_2) \subset (A, B)$. But this is a contradiction since $(A, B)$ is assumed to be the mincut. $\Box$

Lemma 2. Let $(A, B)$ be a mincut of $G$, and let $\lambda(G) = |(A, B)| < \delta(G)$. Then there exists a node $x \in A$, such that $x \notin N(B)$. Similarly $\exists y \in B$ such that $y \notin N(A)$.

**Proof :** Suppose that every node in $A$ is adjacent to some node in $B$. Let $u$ be a node in $A$. Let $F = (A, B)$, be the minimum cut. We have

$$d(u) = |N(u) \cap B| + |N(u) \cap A| \leq |N(u) \cap B| + |A| - 1$$

But

$$|F| = \sum_{x \in A} |N(x) \cap B| \geq |N(u) \cap B| + |A| - 1$$

since each term in the sum should be at least 1 by the assumption that every node in $A$ is adjacent to at least one node in $B$. It follows that $d(u) \leq |F|$. But by assumption, $d(u) \geq \delta(G) > |F|$. Thus we have a contradiction and we conclude that there is a node $x \in A$ such that $x \notin N(B)$. By similar arguments, $\exists y \in B$ such that $y \notin N(A)$. $\Box$

Corollary 1. If $(A, B)$ is a mincut of $G$, and if $\lambda(G) \neq \delta(G)$ then $|A| > \kappa(G)$, $|B| > \kappa(G)$.
Proof: By lemma 2, there exists a node $x \in A$ such that $x \notin N(B)$. Let $y \in B$. It is clear that $N(B)$ separates $x$ from $y$. Thus $|N(B)| \geq \kappa(G)$. But $N(B) \subseteq A - \{x\}$. It follows that $\kappa(G) < |A|$. Similarly $\kappa(G) < |B|$. □

A chordless path from $u$ to $v$ is defined to be a path from $u$ to $v$ in $G$ with no two non consecutive nodes of the path being adjacent. The reader can easily verify that if there is a path between $u$ and $v$ then there is a chordless path also. For example, the shortest path between $u$ and $v$ has to be a chordless path.

Lemma 3. Let $P = (w_1, w_2, \ldots, w_k)$ be a chordless path in a chordal graph $G$ and let $w_i = \text{highest}(P)$. Then $(w_1, w_2, \ldots, w_i)$ is an increasing path while $(w_i, w_{i+1}, \ldots, w_k)$ is a decreasing path.

Proof: Consider first the path $(w_1, w_2, \ldots, w_i)$. Since $w_i = \text{highest}(P)$, this cannot be a decreasing path. If $i \leq 2$, obviously this is an increasing path. Let $i > 2$. If $(w_1, w_2, \ldots, w_i)$ is not an increasing path, then there should be a node $w_j \in P$, $1 < j < i$, such that $\text{PEO}(w_j) < \text{PEO}(w_{j+1})$ and $\text{PEO}(w_j) < \text{PEO}(w_{j-1})$. Then \{w_{j-1}, w_{j+1}\} \subseteq \mathcal{N}_h(w_j).$ By definition of PEO, $\mathcal{N}_h(w_j)$ forms a clique, and we have $(w_{j-1}, w_{j+1}) \in E$. So $(w_{j-1}, w_{j+1})$ forms a chord for $P$, contradicting the chordless property of $P$. We infer that $(w_1, w_2, \ldots, w_i)$ is an increasing path. A similar argument shows that $(w_i, w_{i+1}, \ldots, w_k)$ should be a decreasing path. □

Corollary 2. Let $P = (w_1, w_2, \ldots, w_k)$ be a chordless path and $w_k = \text{highest}(P)$. Then $P$ is an increasing path.

Corollary 3. Let $P = (w_1, w_2, \ldots, w_k)$ be a chordless path and $\text{PEO}(w_1) < \text{PEO}(w_k)$ Then $\text{PEO}(w_2) > \text{PEO}(w_1)$.

Let $G = (V, E)$ be a chordal graph and $(A, B)$ be a mincut of $G$. Without loss of generality assume that $B$ contains the node with PEO number $n$. Let $x = \text{highest}(A)$. Let $B_1 = B \cap N(x)$ and $B_2 = B - B_1$.

Lemma 4. If $|B_1| < \kappa(G)$, and $\lambda(G) \neq \delta(G)$, then $\text{PEO}(\text{highest}(B_2)) < \text{PEO}(\text{highest}(A))$.

Proof: Let $x = \text{highest}(A)$ and $y = \text{highest}(B_2)$. (Note that since $|B| > \kappa(G)$ (by corollary 1), and $|B_1| < \kappa(G)$, $B_2$ is nonempty and therefore $\text{highest}(B_2)$ exists). First observe that $\mathcal{N}_h(x) \subseteq B_1$, since every neighbour of $x$ with a higher PEO number has to be in $B$. We will prove the lemma by contradiction. In other words, we will show that if $\text{PEO}(y) > \text{PEO}(x)$, then $|\mathcal{N}_h(x)| \geq \kappa(G)$, thereby contradicting the assumption that $|B_1| < \kappa(G)$, since $\mathcal{N}_h(x) \subseteq B_1$. Note that $y$ is non adjacent to $x$ since $y \in B_2$. We will show that $\mathcal{N}_h(x)$ separates $x$ from $y$. Consider any path $P$ from $x$ to $y$. Clearly there is a chordless path $P'$ from $x$ to $y$, containing only the nodes in $P$. Let $w$ be the second node in $P'$, starting from $x$. By corollary 3 $\text{PEO}(w) > \text{PEO}(x)$. Thus $w \in \mathcal{N}_h(x)$. We have shown that in every path $P$ from $x$ to $y$ there is a node which is in $\mathcal{N}_h(x)$. Therefore $\mathcal{N}_h(x)$ is a separator and $|\mathcal{N}_h(x)| \geq \kappa(G)$. This in turn implies $|B_1| \geq \kappa(G)$, contradicting the assumption. □
**Theorem 3.** For a chordal graph $G$, if $\lambda(G) \neq \delta(G)$, then $\lambda(G) \geq 2\kappa(G) - 1$.

**Proof:** Let $(A, B)$ be a mincut of $G$. Remember that $|(A, B)| = \lambda(G) < \delta(G)$. Without loss of generality, let $B$ contain the node with PEO number $n$. Let $x = \text{highest}(A)$. By lemma [2], there is a node $a \in A$, and a node $b \in B$ such that $a \notin N(B)$ and $b \notin N(A)$. We observe first that $a \neq x$. To see this, consider a chordless path from $x$ to the node with PEO number $n$, which is in $B$ by assumption. By corollary [2] this is an increasing path and since $x = \text{highest}(A)$, the second node in this path after $x$ has to be in $B$. It follows that $x \in N(B)$. Therefore $x \neq a$.

Now since $a$ and $b$ are nonadjacent, by Menger’s theorem (Theorem 2), there exist at least $\kappa(G)$ vertex disjoint paths between them. One of these paths may contain $x$. Consider $\kappa(G) - 1$ vertex disjoint paths between $a$ and $b$, which do not contain $x$. Each of these paths should exit $A$, thus contributing at least one edge to the mincut $(A, B)$. Collect one such edge from each path and let $F_1$ be the set of $\kappa(G) - 1$ edges collected this way. We emphasize that none of the edges in $F_1$ is incident on $x, a$ or $b$. (Since we have avoided the path containing the node $x$, and none of the edges incident on $a$ or $b$ are in $(A, B)$, by the choice of $a$ and $b$). Moreover no two edges in $F_1$ will share a common end point since they are all coming from distinct vertex disjoint paths.

Now let $B_1 = N(x) \cap B$, be the set of neighbours of $x$ which are in $B$. Let $F_2 = \{(x, u) : u \in B_1\}$, the set of edges starting from $x$ and ending at a node in $B_1$. Clearly $F_2 \subseteq (A, B)$ and $F_2 \cap F_1 = \phi$. If $|B_1| \geq \kappa(G)$, $|F_2| \geq \kappa(G)$ also and we have $\lambda(G) = |(A, B)| \geq |F_1| + |F_2| \geq 2\kappa(G) - 1$, as required by the theorem. If $|B_1| < \kappa(G)$, let $B_2 = B - B_1$. $B_2$ is nonempty since $|B| > \kappa(G)$ by corollary [1]. Let $y = \text{highest}(B_2)$. By lemma [4] $\text{PEO}(y) < \text{PEO}(x)$. Noting that $x$ and $y$ are nonadjacent, there exist $\kappa(G)$ vertex disjoint paths between them, by Menger’s theorem (Theorem 2). Clearly at least $l = \kappa(G) - |B_1|$ of them should be paths which consist of vertices in $A \cup B_2$ only. Let these paths be $P_1, P_2, \ldots, P_l$. For each path $P_i$, $1 \leq i \leq l$, there exists a chordless path $P_i'$ from $y$ to $x$ which uses only the nodes of $P_i$. Since $x$ and $y$ are nonadjacent, each of these chordless paths should contain at least 3 nodes. Let $u_i$ be the second node in $P_i'$, starting from $y$. We claim that $u_i \in A$. This is because $\text{PEO}(u_i) > \text{PEO}(y)$ by corollary [8] and $y = \text{highest}(B_2)$. Since $u_i \notin B_1$, it has to be in $A$. Thus $(u_i, y) \in (A, B)$. Now consider a chordless path in $B$ from $y$ to the node having PEO number $n$. Since $G(B)$ is connected by lemma [1] such a chordless path exists and clearly it is an increasing path. Let $w$ be the second node in this path starting from $y$. Clearly $u_i$ and $w$ are neighbours of $y$ with higher PEO number than that of $y$ itself, ie $\{w, u_i\} \subseteq N_h(y)$. By definition of PEO, $N_h(y)$ forms a clique and therefore $(u_i, w) \in E$. Clearly since $w \in B$ and $u_i \in A$, $(u_i, w) \in (A, B)$. Since no two edges in $F_1$ shares a common end point, one of the two edges $(u_i, y)$ or $(u_i, w)$ is not in $F_1$. Now for each $P_i'$, $1 \leq i \leq l$, add to $F_3$, one of $(u_i, w), (u_i, y)$ which ever is not in $F_1$. Clearly $|F_3| = l - \kappa(G) - |B_1|$. Also $F_3 \cap F_1 = \phi$ and $F_3 \cap F_2 = \phi$. Since $F_1 \cup F_2 \cup F_3 \subseteq (A, B)$, we have $\lambda(G) = |(A, B)| \geq |F_1| + |F_2| + |F_3| = 2\kappa(G) - 1$, as required by the theorem. This completes the proof. □
References

Changing the Diameter of Graph Products

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Abstract. Graham and Harary [3] studied how the diameter of hypercubes can be affected by increasing and decreasing edges. Since many networks are constructed by graph products, e.g., tori and meshes, in this paper we study how the diameter of graph products, particularly Cartesian product and Kronecker product, can be changed by adding and removal of edges. We study Cartesian products on paths, cycles, trees, and hypercubes. The diameter of the Kronecker product of two graphs is in general difficult to find. We in particular study the Kronecker product of two cycles.

1 Introduction

Let $G = (V, E)$ denote a graph. For any two vertices $u, v$ in $G$, let $d_G(u, v)$ denote the distance between $u$ and $v$ which is defined as the length of a shortest path between them. The diameter of $G$, denoted by $D(G)$, is defined as $D(G) = \max_{u,v \in V} d_G(u, v)$. Graham and Harary [3] studied how the diameter of hypercubes can be affected by increasing and decreasing edges. We are interested in the following measures defined in [3]:

$D^-(G)$: the least number of edges whose addition to $G$ decreases the diameter;

$D^+(G)$: the least number of edges whose deletion from $G$ increases the diameter.

For example, since deleting an edge from a cycle of length $m$, denoted by $C_m$, increases the diameter from $\lceil m/2 \rceil$ to $m - 1$, it follows that $D^+(C_m) = 1$. The superscript of $D$ indicates the “change” of diameter, i.e., “-” to mean decreasing the diameter, and “+” increasing the diameter.

Many popular networks are constructed by graph products. We are particularly interested in Cartesian product and Kronecker product. Let $G_1$ and $G_2$ be two graphs, and we write $G_1 = (V_1, E_1)$ and $G_2 = (V_2, E_2)$ throughout the paper. The Cartesian product $G = (V, E)$ of $G_1$ and $G_2$, denoted by $G = G_1 \cdot G_2$, is given by $V = V_1 \times V_2$, and

$$E = \{(u_1u_2, v_1v_2) \mid u_1 = v_1 \text{ and } (u_2, v_2) \in E_2, \text{ or } u_2 = v_2 \text{ and } (u_1, v_1) \in E_1\}.$$ 

The Kronecker product $G = (V, E)$, also known as direct product and tensor product, of $G_1$ and $G_2$, denoted by $G = G_1 \times G_2$, is given by $V = V_1 \times V_2$ and

$$E = \{(u_1u_2, v_1v_2) \mid (u_1, v_1) \in E_1 \text{ and } (u_2, v_2) \in E_2\}.$$
A torus is the Cartesian product of two cycles, and mesh the Cartesian product of two paths. The diagonal toroidal mesh discussed in [6] is the Kronecker product of two cycles. In [2], the authors studied the Cartesian products of complete binary trees, shuffle-exchange graphs and de Bruijn graphs, which possess various good properties.

We adopt the following notation. Let $P_n$ and $C_n$ denote a path and a cycle, respectively, of $n$ vertices labeled by $0, 1, \ldots, n - 1$. Let $Q_n$ be a hypercube of dimension $n$ of $2^n$ vertices with each vertex denoted by an $n$-bit binary representation, where two vertices are connected by an edge if and only if two vertices differ at exactly one bit.

Since we work on graph products, for ease of exposition we use $(\cdot)$, $[\cdot]$ and $(\cdot)$ to delimit a vertex, a pair of vertices and an edge, respectively. For example, each vertex in $G = G_1 \cdot G_2$ is denoted by $(v_1, v_2)$, where $v_1 \in V_1$ and $v_2 \in V_2$. An edge in $G_1 \cdot G_2$ is denoted by $\langle(u_1, u_2), (v_1, v_2)\rangle$. We also use $(\cdot)$ and $(\cdot)$ to delimit a path.

## 2 Preliminaries

Let $\delta(G)$ denote the minimum degree of $G$, i.e., $\delta(G) = \min \{\delta_C(v) \mid v \in G\}$, where $\delta_C(v)$ denotes the degree of a vertex $v$ in $G$. It is obvious that $D^+(G) \leq \delta(G)$.

It is known that $D(P_n) = n - 1$, $D^-(P_n) = 1$ and $D^+(P_n) = 1$. The diameter of cycle $C_n$ is given by $D(C_n) = \lceil n/2 \rceil$. Obviously, $D^+(C_n) = 1$. It is shown in [1] that $D^-(C_n) = 2$ for $n \geq 8$ by adding the following two edges:

$$
e_1 = \langle 0, \lceil n/2 \rceil \rangle,
$$

$$
e_2 = \begin{cases} 
\langle \lceil n/4 \rceil, \lfloor 3n/4 \rfloor \rangle & \text{for } n = 2 \mod 4, \\
\langle \lceil n/4 \rceil, \lfloor 3n/4 \rfloor \rangle & \text{for } n = 0, 1, 3 \mod 4.
\end{cases} \tag{1}
$$

Hypercubes have been extensively studied. Let $v$ be a vertex in $Q_n$. It is known that $Q_n$ is n-connected. Let $x$ and $y$ be two vertices in $Q_n$. We use $h[x, y]$ to denote the hamming distance between $x$ and $y$ which is defined as the number of differing bits in $x$ and $y$. It is shown in [3] that $D^-(Q_n) = 2$ by adding two diagonal edges in any subcube $Q_2$ of $Q_n$, e.g., $\langle v_0, 00\ldots 011 \rangle$ and $\langle 00\ldots 001, 00\ldots 010 \rangle$. It is well known and shown in [5] that there are $n$ internally node-disjoint paths between $x$ and $y$, $h[x, y]$ of them having length of $h[x, y]$ and $n - h[x, y]$ having length of $h[x, y] + 2$. It follows that $D^+(Q_n) = n - 1$ by deleting $n - 1$ edges in a subcube $Q_{n-1}$ of $Q_n$ which are incident at any specific vertex $[3]$. For example, deleting $\langle 000, 001 \rangle$ and $\langle 000, 010 \rangle$ from $Q_3$, the distance from 000 to 011 becomes 4, so is the diameter.

## 3 Cartesian Products

Cartesian products also preserve the node symmetry property of its two primitive graphs. To be specific, let $G_1, G_2$ be node-symmetric. The following lemma can be easily proved.
Lemma 1 Let $G_1, G_2$ be node-symmetric. Then $G_1 \cdot G_2$ is also node-symmetric.

The following lemma can be easily shown.

Lemma 2 $D(G_1 \cdot G_2) = D(G_1) + D(G_2)$.

For example, $D(T_{m,n}) = D(C_m) + D(C_n) = \lfloor m/2 \rfloor + \lfloor n/2 \rfloor$. Furthermore, it can be easily verified that

$$D^+(G_1 \cdot G_2) \geq \min \{D^+(G_1), D^+(G_2)\} + 1 \quad (2)$$

$$D^-(G_1 \cdot G_2) \geq \min \{D^-(G_1), D^-(G_2)\} \quad (3)$$

One may ask whether (2) and (3) always hold as inequalities. To answer this, consider $P_m \cdot P_n$ which is called mesh of size $m$ by $n$. It can be easily verified that $D^-(P_m \cdot P_n) = 1$ and $D^+(P_m \cdot P_n) = 2$, i.e., (2) and (3) hold as equalities. We subsequently ask how to characterize those $G_1$ and $G_2$ for (2) and (3) to hold, respectively, as equalities. Another interesting question is whether we can tighten the lower bound in (2).

On the other hand, we wonder whether (2) and (3) can hold as equalities when $G_1$ and $G_2$ are node-symmetric. Consider $T_{m,n} = C_m \cdot C_n$, where $C_i$ is node-symmetric. We can show the following theorem with proof omitted.

Theorem 1 $D^+(C_m \cdot C_n) = \begin{cases} 2 & \text{for } m, n \text{ odd}, \\ 3 & \text{for one of } m, n \text{ odd}, \\ 4 & \text{for } m, n \text{ even}. \end{cases}$

The above theorem illustrates that (2) does not necessarily hold as equality when $G_1$ and $G_2$ are node-symmetric. It is shown in (4) that $D^-(C_n) = 2$ for $n \geq 8$ and $D^-(T_{m,n}) = 2$ for $m \geq 12$ or $n \geq 12$, and (3) holds as equality.

In what follows, we study $D^+(G_1 \cdot G_2)$ and $D^-(G_1 \cdot G_2)$ for $G_1, G_2$ to be paths, cycles, hypercubes and trees.

Lemma 3 (i) $D^-(C_m \cdot P_n) = 1$ for all $n \geq 4$.
(ii) $D^+(C_m \cdot P_n) = \begin{cases} 2 & \text{for } m \text{ odd}, \\ 3 & \text{for } m \text{ even}. \end{cases}$
(iii) $D^-(Q_m \cdot P_n) = 1$ for $n \geq 3$.
(iv) $D^+(Q_m \cdot P_n) = m$.
(v) $D^-(Q_m \cdot C_n) = 2$ for $n \geq 8$.
(vi) $D^+(Q_m \cdot C_n) = \begin{cases} m + 2 & \text{for } n \text{ even}, \\ m + 1 & \text{for } n \text{ odd}. \end{cases}$
(vii) $D^-(T_1 \cdot T_2) = \min \{D^-(T_1), D^-(T_2)\}$ and $D^+(T_1 \cdot T_2) = 2$, where $T_1$ and $T_2$ are arbitrary trees.

After careful observation of the above theorem, we have the following conjecture.

Conjecture 1 If $G_1$ and $G_2$ are node-symmetric, then

$$D^-(G_1 \cdot G_2) = \min \{D^-(G_1), D^-(G_2)\}.$$  

Furthermore, we conjecture that the lower bound in (2) can be increased as stated below.

Conjecture 2 $D^+(G_1 \cdot G_2) \geq \max \{D^+(G_1), D^+(G_2)\} + 1.$
4 Kronecker Product of Cycles

Kronecker product preserves good properties of its two primitive graphs. For example, if $G_1$ and $G_2$ are node-symmetric, $G_1 \times G_2$ is also node-symmetric.

The diameter of a Kronecker product on two arbitrary graphs is in general difficult to find. In this section, we consider Kronecker products of two cycles $C_m$ and $C_n$. In this section, we calculate $D(C_m \times C_n)$ using different approach from that in [6] and find $D^+(C_m \times C_n)$ and give an upper bound for $D^-(C_m \times C_n)$. If both $m$ and $n$ are even, $C_m \times C_n$ becomes disconnected. In this section, we assume that at least one of $m$ and $n$ is odd.

In $C_k$, let $N(i)$ denote the $i$th-neighborhood of vertex 0 which is defined as the set of the vertices in $C_k$ that can be reached from 0 at the $i$th step. Then

$$N(i) = \begin{cases} \{1, 3, \ldots, i, k - i, \ldots, k - 1\} & \text{for } i \text{ odd}, \\ \{0, 2, 4, \ldots, i, k - i, k - i + 2, \ldots, k - 2\} & \text{for } i \text{ even}. \end{cases}$$

For example, given $k = 11$, we have

$$N(1) = \{1, 10\}, \quad N(2) = \{0, 2, 9\}, \quad N(5) = \{1, 3, 5, 6, 8, 10\},$$

$$N(6) = \{0, 2, 4, 5, 6, 7, 9\}, \quad N(9) = \{1, 2, 3, 4, \ldots, 10\}.$$  

With careful observation, we can obtain the following remark.

Remark 1 (i) $i = \lfloor k/2 \rfloor$ is the smallest integer satisfying

$$\bigcup_{j=1}^i N(j) = \{0, 1, 2, \ldots, k - 1\},$$

and in particular, $N(\lfloor k/2 \rfloor - 1) \cup N(\lfloor k/2 \rfloor) = \{0, 1, 2, \ldots, k - 1\}$.

(ii) $\lfloor k/2 \rfloor \notin N(i)$ for all $i \leq \lfloor k/2 \rfloor - 1$.

(iii) $N(i) \subseteq N(i + 2)$ for $1 \leq i \leq n - 3$, and in particular, $N(i) = N(i + 2)$ for $k$ even and $k/2 \leq i \leq k - 3$.

In this section, let $G = C_m \times C_n$ and $d^* = D(G)$. Since $G$ is node-symmetric, we can consider $d^*$ as the longest distance from $(0, 0)$ to all of the other vertices. In order to distinguish the $i$th-neighborhood of vertex 0 in $C_m$ and $C_n$, we use $N_1(i)$ and $N_2(i)$ to denote that in $C_m$ and $C_n$, respectively. Henceforth, we use $N(i)$ to denote the $i$th-neighborhood of vertex $(0, 0)$ in $G$ without ambiguity, i.e., the set of vertices in $G$ that can be reached from $(0, 0)$ at the $i$th step. It follows that $N(i) = N_1(i) \times N_2(i)$. The diameter $d^*$ is the smallest integer $i$ satisfying $\bigcup_{j=1}^i N(j) = V_1 \times V_2$. It can be easily verified that $d^* \geq \max\{\lfloor m/2 \rfloor, \lfloor n/2 \rfloor\}$.

In this paper we consider that both $m$ and $n$ are odd and assume without loss of generality that $n \geq m$. The case with exactly one of $m$ and $n$ being odd can be similarly treated. We have the following results.

Lemma 4 Let both $m$ and $n$ be odd, and $n \geq m$. Then

$$D(C_m \times C_n) = \begin{cases} m - 1 & \text{for } m = n, \\ m & \text{for } m + 2 \leq n \leq 2m + 1, \\ \lfloor n/2 \rfloor & \text{for } n \geq 2m + 3. \end{cases}$$
Proof: We assume without loss of generality that $m \leq n$. To find $d^*$, we distinguish the following cases:

Case (i): $m = n$.
Since $N_1(m-1) = V_1$ and $N_1(i) = N_2(i)$, it follows that $N(m-1) = V_1 \times V_2$ and $d^* \leq m-1$. On the other hand, since $N_1(m-1) = N_1(m-2) = \{0, m-1\}$ and $N_1(m-2) - N_1(m-3) = \{1\}$, it follows that the vertices $(0,1), (1,0), (0,m-1)$, and $(m-1,0)$ are not in $N(m-2)$. Therefore, $m-2 < d^* \leq m-1$, i.e., $d^* = m-1$.

Case (ii): $m + 2 \leq n \leq 2m + 1$.
Since $m \geq \lfloor n/2 \rfloor$, it follows that $N_2(m-1) \cup N_2(m) = V_2$. Furthermore, since $N_1(m) = V_1$, it follows that $N(m) = V_1 \times V_2$ and thus $d^* \leq m$. On the other hand, since $N_1(m-1) - N_1(m-2) = \{0\}$, $N(m-1)$ does not contain the vertex set $\{(0,j) \mid j \in N_2(m) - N_2(m-1)\}$. It follows that $m-1 < d^* \leq m$, i.e., $d^* = m$.

Case (iii): $n \geq 2m + 3$, i.e., $\lfloor n/2 \rfloor \geq m + 1$.
Note that $N_1(\lfloor n/2 \rfloor - 1) = N_1(\lfloor n/2 \rfloor) = V_1$ and $d^* \geq \lfloor n/2 \rfloor$. On the other hand, since $N_2(\lfloor n/2 \rfloor - 1) \cup N_2(\lfloor n/2 \rfloor) = V_2$ as stated in Remark II(i), it follows that $\bigcup_{i=1}^{\lfloor n/2 \rfloor} N_1(i) \times N_2(i) = V$. Thus $d^* = \lfloor n/2 \rfloor$. The lemma follows.

Our proof method for Lemma 4 is different from that in [6]. We then can show the following theorem with proof omitted.

Theorem 2 Let $m, n$ be odd, and $n \geq m$.

$$D^+(C_m \times C_n) = \begin{cases} 1 & \text{if } n = 2m + 1, \\ 2 & \text{otherwise.} \end{cases}$$

We can also give a construction of adding edges such that the diameter is reduced which also provides an upper bound of $D^-(C_m \times C_n)$.

References

Plane Graphs with Acyclic Complex*

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Abstract. All vertex-sets of cliques of \( G \) form the simplexes of an abstract complex \( C(G) \), which is called the *clique complex* of \( G \). It is proved that the clique complex of a plane graph \( G \) is acyclic if and only if \( G \) is contractible.

1 Introduction

Graphs considered are finite and simple, complexes considered are finite. Undefined signs can be found in [1,2]. An \( n \)-complex is a complex of dimension \( n \). \( n \)-simplex, \( n \)-chain and \( n \)-cycle are defined analogously. The topology properties of graphs have important applications in image analyses [3].

Given an \( n \)-complex \( K \) and integer \( 0 \leq q \leq n \), \( H_q(K) \) and \( \chi(K) \) denote the \( q \)-th homology group of \( K \) with coefficients in \( \mathbb{Z} \) and the Euler characteristic of \( K \), respectively. \( \partial \) is the boundary operator on \( K \). An \( n \)-complex \( K \) is called acyclic if \( H_q(K) = 0 \) for \( q > 0 \). Two complexes \( K \) and \( L \) are homologous, denoted by \( K \sim L \), if \( H_q(K) = H_q(L) \) for each \( q \geq 0 \).

Given a graph \( G = (V(G), E(G)) \), \( N_G(u) \) denotes the set of vertices adjacent to \( u \) in \( G \). \( d_G(u) \) and \( \delta(G) \) denote the degree of \( u \) and the minimum degree of \( G \), respectively. A \( k \)-vertex is vertex of degree \( k \). For a set \( S \subset V(G) \), \( G[S] \) denotes the subgraph of \( G \) induced by \( S \). A *clique* is a complete subgraph of \( G \). All vertex-sets of cliques of \( G \) form the simplexes of an abstract complex \( C(G) \), which is called the *clique complex* of \( G \). For convenient, we use \( H_q(G) \) to denote the \( q \)-th homology group of \( C(G) \), and use a sequence of vertices of \( G \) to denote an oriented simplex of \( C(G) \).

**Definition 1** [4] (Ivashchenko) A family \( \mathcal{R} \) of graphs is called contractible if \( K_1 \in \mathcal{R} \) and each graph in \( \mathcal{R} \) can be obtained from \( K_1 \) by following contractible transformations: (1) Deleting a vertex \( v \) from \( G \), where \( G[N_G(v)] \in \mathcal{R} \); (2) Adding a vertex \( v \notin V(G) \) to \( G \) and joining it to each vertex of an induced subgraph \( G_1 \in \mathcal{R} \) of \( G \); (3) Deleting an edge \( uv \in E(G) \), where \( G[N_G(u) \cap N_G(v)] \in \mathcal{R} \); (4) Adding an edge \( uv \notin E(G) \) to \( G \), where \( G[N_G(u) \cap N_G(v)] \in \mathcal{R} \). All graphs in \( \mathcal{R} \) are called contractible.

**Theorem 1** [4] (Ivashchenko) Contractible transformations do not change the homology groups of graphs. Especially, if a graph \( G \) is contractible, then \( H_0(C(G)) = J \), and \( H_q(C(G)) = 0, q \geq 1 \).

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By Theorem 11, \( C(G) \) is acyclic while \( G \) is contractible. We guess that each connected graph with acyclic clique complex is contractible. Our main result is

**Theorem 12** Let \( G \) be a connected plane graph. Then \( C(G) \) is acyclic if and only if \( G \) is contractible.

## 2 Homology Groups of Plane Graphs

Let \( G \) be a plane graph. When we say that \( H \) is a subgraph of \( G \), it means that the plane embedding of \( H \) is a restriction of \( G \) on \( H \). For a face \( f \) of \( G \), we use \( b(f) \) to denote a sequence of vertices in clockwise order on the boundary of \( f \). A face \( f \) of \( G \) is called a \( k \)-face if \( f \) is incident with exactly \( k \) edges. A maximal subgraph of \( G \) is a maximal plane graph which is a subgraph of \( G \). Let \( C \) be a cycle of \( G \), we use \( Int_G(C) \) to denote the set of vertices located in the interior of \( C \).

Let \( H \) be a maximal subgraph of a plane graph, \( \{f_1, f_2, \ldots, f_i\} \) the set of faces of \( H \). Suppose that \( f_1 \) is the unbounded face of \( H \). We define \( link(H) = \sum_{i=2}^i b(f_i) - b(f_1) \). Then \( link(H) \) is a cycle of \( C(G) \).

We use \( T(G) \) to denote the set of maximal subgraphs of \( G \) which has at least four vertices and contains no separating triangles (a triangle \( uvw \) is a separating triangle if both the interior and the exterior of \( uvw \) contain vertices of \( G \)), and let \( LT(G) = \{link(H) | H \in T(G)\} \).

One can easily check the following

**Lemma 21** If \( G \) is a plane graph, then \( H_q(G) = 0 \) for \( q \geq 3 \).

**Lemma 22** If \( G \) is a plane graph, then any three distinct 3-simplexes of \( C(G) \) can not share a common 2-simplex.

Let \( x = \sum_{i=1}^{s_q} r_i s_i^q (r_i \in J) \) and \( z = \sum_{i=1}^{s_q} s_i^q (s_i \in J) \) be two \( q \)-cycles of a complex \( K \). If \( |s_i| \leq |r_i| (i = 1, 2, \ldots, \alpha_q) \), \( r_i s_i > 0 \) while \( s_i \neq 0 \), and \( r_{i_0} \neq s_{i_0} \) for some \( i_0 \in \{1, 2, \ldots, \alpha_q\} \), then \( z \) is called a proper subcycle of \( x \). A basic cycle is a cycle containing no proper subcycle.

**Lemma 23** Let \( G \) be a plane graph, \( c \) a 2-cycle of \( C(G) \). Then, \( c \) is a basic cycle if and only if \( c = link(H) \) for some maximal plane subgraphs \( H \neq K_3 \).

**Proof.** The sufficiency is clearly. Following we prove the necessity by contradiction. If it is not so, let \( c \) be a basic 2-cycle and \( c \neq link(H) \) for any maximal plane subgraph \( H \neq K_3 \). Let \( H_0 \) denote the subgraph of \( G \) induced by the edges corresponding to the 1-simplexes in \( c \). If \( H_0 \) is not a maximal plane graph, then it must has a non-triangular face, say \( f_0 \), let \( u_0v_0 \) be an edge incident with \( f_0 \).

Without loss of generality, assume that \( f_0 \) is the unbounded face of \( H_0 \). Because \( c \) is a cycle, there must be two distinct triangles \( x_0u_0v_0x_0 \) and \( y_0u_0v_0y_0 \) in \( H_0 \) such that \( x_0 \in \text{Int}_{H_0}(y_0u_0v_0) \). Let \( G_0 = H_0[\{y_0, u_0, v_0\} \cup \text{Int}_{H_0}(y_0u_0v_0y_0)] \). Then \( |V(G_0)| < |V(G)| \).
If there still exist \( l \)-faces in \( G_0 \) with \( l \geq 4 \), then take such a face, say \( f_1 \), and an edge \( u_1v_1 \) incident with \( f_1 \). There must be two distinct triangles \( x_1u_1v_1x_1 \) and \( y_1u_1v_1x_1 \) in \( G_0 \) such that \( x_1 \in \text{Int}_{G_0}(y_1u_1v_1) \) because \( c \) is a cycle. Let \( G_1 = G_0[[y_1, u_1, v_1] \cup \text{Int}_{G_0}(y_1u_1v_1)] \). Then, \( |V(G_1)| < |V(G_0)| \).

Go ahead so on and so forth. Because \( V(H_0) \) is finite, there must be a triangle \( yuvy \) in \( H_0 \) such that \( \text{Int}_{H_0}(yuvy) \neq \emptyset \) and all faces of \( H_0 \) in the interior of \( yuvy \) are 3-faces. Let \( H' = H_0[[y, u, v] \cup \text{Int}_H(yuvy)] \), and let \( c' = \text{link}(H') \). Clearly, \( c' \) is a proper subcycle of \( c \). A contradiction.

Let \( G_1 \) and \( G_2 \) be two plane graphs, \( f_1 \) with boundary \( u_1u_2u_3u_4 \) and \( f_2 \) with boundary \( v_1v_2v_3v_4 \) be two 3-faces of \( G_1 \) and \( G_2 \), respectively. We can embed \( G_2 \) into the 3-face \( f_1 \) by identifying \( u_i \) with \( v_i \) (\( i = 1, 2, 3 \)), respectively. This new plane graph is called a triangle-embedding of \( G_2 \) in \( G_1 \).

**Lemma 24** Let \( G \) be a plane graph, \( c \) a basic 2-cycle of \( C(G) \). Then,

\[
c \sim \sum_{c \in \text{LT}(G)} r(c)c, \text{ where } r(c) = 0, \text{ or } 1.
\]

**Proof.** By Lemma 23, there exists a maximal subgraph \( H \neq K_3 \) such that \( c = \text{link}(H) \). Without loss of generality, suppose \( H \not\supseteq K_4 \) and \( H \not\supseteq T(G) \), \( u, w, v \) are the three vertices incident with the unbounded face of \( H \) in clockwise order. The proof is by induction on the order of \( H \). It is obviously that \( |V(H)| \geq 5 \). While \( |V(H)| = 5 \), \( H \) is isomorphic to the graph obtained by embedded \( K_4 \) in \( K_4 \). Obviously, \((1)\) holds.

Assume that \((1)\) is valid while \( |V(H)| \leq n \) (\( n \geq 5 \)). Following suppose \( |V(H)| = n + 1 \).

**Case 1.** There exists \( x \in \text{Int}_H(uuvw) \) such that \( H[\{u, v, w, x\}] \cong K_4 \). Since \( H \not\supseteq K_4 \), then at least one of the three cycles \( xwv, xuvx \) and \( xuwx \) is nonempty. Without loss of generality, suppose \( xuw, xuvx, xuvx \) are all nonempty.

Let \( c_{uw} = \text{link}(H[\{x, u, v\} \cup \text{Int}_H(xwv)]) \), \( c_{uv} = \text{link}(H[\{x, v, w\} \cup \text{Int}_H(xwv)]) \) and \( c_{uw} = \text{link}(H[\{x, u, w\} \cup \text{Int}_H(xwv)]) \). Suppose \( abc \) is a simplex of \( c, q, q_{uw}, q_{uw}, q_{uw} \) denote the coefficients of \( abc \) in \( c, c_{uv}, c_{uv}, c_{uw}, c_{uw} \), respectively. Then, by the definition of \( \text{link}(H) \),

\[
q = q_{uw} = 1, q_{uw} = q_{uw} = 0, \text{ while } \{a, b, c\} \subset \text{Int}_H(xwv), \quad (2)
\]

\[
q = q_{uv} = 1, q_{uw} = q_{uw} = 0, \text{ while } \{a, b, c\} \subset \text{Int}_H(xwv), \quad (3)
\]

\[
q = q_{uw} = 1, q_{uw} = q_{uw} = 0, \text{ while } \{a, b, c\} \subset \text{Int}_H(xwv), \quad (4)
\]

Since 2-simplex \( uuv \) appears only in \( c \), \( uvx \) appears only in \( c_{uv} \), \( xuv \) only in \( c_{uw} \), and \( xuw \) appears only in \( c_{uw} \). Combining \((2)\), \((3)\) and \((4)\), we have

\[
c - c_{uw} - c_{uw} = -uvw + xwv + xuv = \partial uvw, \quad \text{i.e., } c \sim c_{uw} + c_{uw} + c_{uw}. \quad (i)
\]

**Case 2.** There exists separating triangle \( abca \in H \). Let \( H_0 = H[\{a, b, c\} \cup \text{Int}_H(abca)], H_1 = H[V(H) - \text{Int}_H(abca)] \). Then, \( c = \text{link}(H_0) + \text{link}(H_1) \), hence \((i)\) holds by the induction hypothesis.
Lemma 25 Let $G$ be a plane graph. Then chains in $LT(G)$ are not boundary chains and any two chains in $LT(G)$ are not homologous.

Proof: By contradiction. Let $H \in T(G)$, $c = \text{link}(H) = \partial z$ where $z$ is a 3-chain of $C(G)$. Without loss of generality, suppose $x_0u_0v_0w_0$ is a simplex of $z$ such that $x_0 \in \text{Int}_G(u_0v_0w_0u_0)$ and

$$|\text{Int}_G(u_0v_0w_0u_0)| = \min\{|\text{Int}(uvwu)|\exists u \in \text{Int}_G(uvww), xuvww \text{ is a simplex of } z\}.$$

By the definition of $T(G)$, $H$ contains no separating triangle, so $x_0u_0v_0$ is not a simplex of $c$, and hence there exists $y_0 \in \text{Int}_G(x_0u_0v_0x_0) - V(H)$ such that $x_0u_0v_0y_0$ is a simplex of $z$. It is clearly that $|\text{Int}_G(x_0u_0v_0x_0)| < |\text{Int}_G(u_0v_0w_0u_0)|$, contradicts to the choice of $u_0v_0w_0u_0$.

Lemma 26 Let $G$ be a connected plane graph. Then, $H_0(G) = J$, $H_1(G) = [(m + 1) - \chi(C(G))]J$, $H_2(G) = mJ$, where $m = |T(G)|$.

Proof: Because $G$ is connected, then $C(G)$ is connected, by Lemma 24, Lemma 25 $H_0(G) = J$, and $H_2(G) = mJ$. Because each basic 1-cycle of $C(G)$ corresponds to a unique cycle of $G$, then $C(G)$ has no 1 dimensional torsion coefficient, and so, $H_1(G) = [(m + 1) - \chi(C(G))]J$.

A pseudo-maximal plane graph is a 2-connected plane graph which contains at most one non-triangular face.

Lemma 27 Let $G$ be a 2-connected plane graph such that $H_1(G) = 0$, and $f$ a non-triangular face of $G$. Then, there exists a pseudo-maximal subgraph $H$ of $G$ such that $f$ is the unique non-triangular face of $H$.

Proof: Let $C = v_1v_2 \ldots v_nv_1 (n \geq 4)$ be the boundary of $f$, $c = \sum_{i=1}^{n} v_iv_{i+1}$ (mod $n$).

Let $\{s_i^2|i = 1, 2, \ldots, \alpha_2\}$ be the set of 2-simplexes of $C(G)$. Since $H_1(G) = 0$, then there exists a 2-chain of $C(G)$, say $x = \sum_{i=1}^{\alpha_2} a_is_i^2$, such that $c = \partial x$. We chose such a 2-chain with the property that

$$\sum_{i=1}^{\alpha_2} |a_i| = \min\{|\sum_{i=1}^{\alpha_2} b_i||c = \partial \sum_{i=1}^{\alpha_2} b_is_i^2\}.$$  \hspace{1cm} (5)

Let $H$ be the subgraph induced by those edges corresponding to the 1-simplexes contained in $x$. Let $H^* = H + \{uv_i\}_{i=1}^{n}$, $G^* = G + \{uv_i\}_{i=1}^{n}$, where $u \notin V(G)$.

Clearly, $G^*$ is a plane graph, $H^*$ is a subgraph of $G^*$, and $y = x + \sum_{i=1}^{n} uv_{i+1}v_i$ (mod $n$) is a 2-cycle of $C(G^*)$. According to (5), $y$ has no proper subcycles, i.e., $y$ is a basic cycle. By lemma 23, $H^*$ is a maximal plane graph, therefore $f$ is the unique non-triangular face of $H$.

Proof of Theorem 12: We only need to prove the necessity. By definition, we can assume that $G$ is 2-connected. By Lemma 26 and Lemma 27, $|T(G)| = 0$
and $G$ can be obtained by a serial of triangle-embeddings of its pseudo-maximal plane subgraphs. Since contractible transformations are reversible, it suffices to show that each pseudo-maximal plane subgraph of $G$ with acyclic clique complex is contractible.

Let $B$ be a pseudo-maximal plane subgraph of $G$ with acyclic clique complex. Without loss of generality, suppose that $B \not\cong K_3$. It is easy to verify that $B$ is contractible while $|E(B)| \leq 8$. Now we proceed to the induction process.

Suppose each pseudo-maximal plane graph with less than $n$ edges and with acyclic clique complex is contractible, where $n \geq 9$. Let $B$ be a pseudo-maximal plane subgraph of $G$ with acyclic clique complex, $|E(B)| = n$, $C$ the boundary of the non-triangular face of $B$, and $uv$ an edge of $C$. Then we have

$$|N_B(u) \cap N_B(v)| = 1. \tag{6}$$

If it is not so, let $x_1, x_2 \in N_B(u) \cap N_B(v)$ and $Int_B(x_1ux_1) \cap V(B) = \emptyset$.

**Case 1.** $x_1x_2 \in E(B)$. Clearly, $G[[x_1, x_2, u, v]] \cong K_4$. If $Int_G(x_1x_2ux_1) \cup Int_B(x_1x_2ux_1) = \emptyset$, then $d_B(x_1) = d_G(x_1) = 3$. By the definition of contractible transformations, $x_1$ can be deleted from $B$, and by the induction hypothesis, $B$ is contractible. So we assume $Int_G(x_1x_2ux_1) \cup Int_B(x_1x_2ux_1) \neq \emptyset$. Go on and go forth, one can always find a vertex $x_0 \in Int_G(ux_2u)$ such $d_G(x) = 3$ and $\{x\} \cup N_B(x)$ induced a $K_4$, then $x_0$ can be deleted from $B$, and so $B$ is contractible by the induction hypothesis.

**Case 2.** $x_1x_2 \notin E(B)$. There must be a vertex in $Int_B(ux_1vx_2)$ which is adjacent to $x_1$ because $B$ is a pseudo-maximal plane subgraph. Therefore, either there exists a 3-vertex $a$ in $B$ such that $B[\{a\} \cup N_B(a)] \cong K_4$ which indicates that deletion of $a$ from $B$ is a contractible transformation, or $B$ contains a subgraph which is in $T(G)$ which contradicts to that $C(G)$ is acyclic.

Now, (6) is proved. Deletion of $uv$ from $B$ is a contractible transformation. Let $B_1 = B - uv$. If $\delta(B_1) = 1$, then one of $u$ or $v$, say $u$, is a 2-vertex in $B$, so deletion of $u$ from $B$ is a contractible transformation, and hence by the induction hypothesis, $B$ is contractible. If $\delta(B_1) \geq 2$, then $B_1$ is a pseudo-maximal plane subgraph of $B$ and $|E(B_1)| = n - 1$, again by the induction hypothesis, $B_1$ is contractible.

**References**

On the Domination Numbers of Generalized de Bruijn Digraphs and Generalized Kautz Digraphs (Extended Abstract)

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Abstract. This work deals with the domination numbers of generalized de Bruijn digraphs and generalized Kautz digraphs. Dominating sets for digraphs are not familiar compared with dominating set for undirected graphs. Whereas dominating sets for digraphs have more applications than undirected graphs. We construct dominating sets of generalized de Bruijn digraphs under some conditions. We investigate consecutive minimum dominating set of the generalized de Bruijn digraphs. For generalized Kautz digraphs, there is a consecutive minimum dominating set that is a minimum dominating set.

1 Introduction

It is one of major areas in theoretical and algorithmic observation to study domination and its related topics for undirected graphs [1,11]. Besides domination has been defined in digraphs, concepts of solution and kernel which are peculiar to digraphs have been defined. The origin of solution and kernel is in game theory [10]. Solution and kernel have applications for logic and facility location [17,9,16]. Considering applications of domination and its related topics for game theory and logic, digraphs are more natural and have more applications than undirected graphs. Although digraphs have many applications, there are few studies for domination in digraphs. Barkauskas and Host, and Bar-Yehuda and Vishkin [2,3] have studied algorithms for domination of digraphs. Barkauskas and Host showed that the problem of determining whether or not an arbitrary digraph has an efficient dominating set is NP-complete and gave a linear time algorithm for finding efficient dominating sets in directed trees [2]. There are some researches for kernel in the digraph with no odd circuits [6] and for domination in tournament that is a class of digraphs [8,14].

Each edge has orientation in digraphs, these edges are called arcs. This situation cause difficulty of research for domination in digraphs.
2 Definitions and Terminologies

$V(G)$ and $A(G)$ are the vertex set and the arc set of a digraph $G(V, A)$, respectively. There is an arc from $x$ to $y$ if $(x, y) \in A(G)$. The vertex $x$ is called a predecessor of $y$ and $y$ is called a successor of $x$. The sets $O(u) = \{v|(u, v) \in A(G)\}$ and $I(u) = \{v|(v, u) \in A(G)\}$ are called the outset and the inset of the vertex $u$, respectively. Similarly, we define $O[u] = O(u) \cup \{u\}$, $I[u] = I(u) \cup \{u\}$. If $S \subset V(G)$ then $O(S) = \cup_{s \in S}O(s)$, $I(S) = \cup_{s \in S}I(s)$, $O[S] = \cup_{s \in S}O[s]$, and $I[S] = \cup_{s \in S}I[s]$. A set $S \subset V(G)$ is a dominating set of a digraph $G$ if $v$ is a successor of some vertex $u \in S$ for all $v \notin S$. A set $S \subset V(G)$ is an absorbant of a digraph $G$ if there is a vertex $u \in S$ such that $u$ is a successor of $v$ for all $v \notin S$. A set $S \subset V(G)$ is independent if for any $x, y \in S, (x, y) \notin A(G)$. If $S \subset V(G)$ is independent and a dominating set of a digraph $G$, $S$ is called a solution of $G$. If $S \subset V(G)$ is independent and an absorbant of a digraph $G$, $S$ is called a kernel of $G$ [10]. $|X|$ denotes the cardinality of a set $X$. For a vertex $u \in V(G)$, $|O(u)|$ is called the out degree of $u$. We use the notation $\gamma(G)$ for the minimum cardinality of a set $S \subset V(G)$ which is a dominating set. If $S$ is a dominating set of $G$ whose cardinality is equal to $\gamma(G)$, $S$ is called the minimum dominating set of $G$. Let $m, n$ be positive integers. $(m, n)$ denotes the greatest common divisor of $m$ and $n$. $m|n$ and $m \not{n}$ mean that $m$ divides $n$ and $m$ does not divide $n$, respectively.

In this paper, we will describe domination numbers of generalized de Bruijn digraphs and generalized Kautz digraphs. Furthermore we construct minimum dominating sets of generalized de Bruijn digraphs and generalized Kautz digraphs. Definitions of generalized de Bruijn digraphs and generalized Kautz digraphs are as follows. Generalized de Bruijn digraph $G_B(n, d)$ is defined by congruence equations.

$$\begin{align*}
V(G_B(n, d)) &= \{0, 1, 2, \ldots, n - 1\}, \\
A(G_B(n, d)) &= \{(x, y)|y \equiv dx + i \pmod{n}, 0 \leq i < d\}.
\end{align*}$$

If $n = d^D$, $G_B(n, d)$ is the de Bruijn digraph $B(d, D)$. Whereas definition of generalized Kautz digraphs $G_K(n, d)$ is

$$\begin{align*}
V(G_K(n, d)) &= \{0, 1, 2, \ldots, n - 1\}, \\
A(G_K(n, d)) &= \{(x, y)|y \equiv -dx - i \pmod{n}, 0 < i \leq d\}.
\end{align*}$$

If $n = d(d - 1)^{D-1}$, $G_K(n, d)$ is the Kautz digraph $K(d - 1, D)$. Generalized de Bruijn digraphs and generalized Kautz digraphs are introduced by Imase and Itoh [12,13], and Reddy, Pradhan and J. Kuhl [15], so called Imase Itoh digraphs.

It is well known that de Bruijn digraphs and Kautz digraphs have good property for network topology compared with hypercubes [5]. These generalizations remove the restriction for the cardinality of vertex sets of these digraphs and make these digraphs more valuable for network models.
3 Domination Numbers of Generalized de Bruijn Digraphs

We prepare a simple lemma for determining domination number of generalized de Bruijn digraphs.

**Lemma 1.** \( \gamma(G_B(n,d)) \geq \left\lceil \frac{n}{d+1} \right\rceil \).

**Proof.** Let \( S \) be a minimum dominating set of \( G_B(n,d) \). We obtain \( |S|+d|S| \geq n \) from the definition of \( G_B(n,d) \).

Above lemma gives the lower bound for the domination number of \( G_B(n,d) \). We also obtain the following lemma concerning the upper bound of the domination number of \( G_B(n,d) \).

**Lemma 2.** \( \gamma(G_B(n,d)) \leq \left\lfloor \frac{n+1}{d} \right\rfloor \).

**Proof.** For \( G_B(n,d) \), let \( S = \{1, 2, \ldots, \lfloor (n+1)/d \rfloor \} \). Then \( O(S) = \{d, d+1, d+2, \ldots, \lfloor (n+1)/d \rfloor d+d-1 \} \) with duplications. \( |O(S)| = \lfloor (n+1)/d \rfloor d+d-1 = \lfloor (n+1)/d \rfloor - 1 \) \( d+d-1 \geq n \). Thus \( S \) is a dominating set of \( G_B(n,d) \). \( \square \)

From the proof of Lemma 2 \( S = \{1,2,\ldots,\lfloor (n+1)/d \rfloor \} \) is a dominating set, but not minimum. For example, the out set of a subset \( \{1,2,3,4\} \) of \( G_B(12,2) \) is \( \{2,3,4,5,6,7,8,9\} \). Then \( \{1,2,3,4\} \) is not a dominating set of \( G_B(10,4) \). Whereas \( \{1,2,3,4,5,6\} \) is a dominating set of \( G_B(10,4) \), still \( \gamma(G_B(12,2)) = 4 \) because the vertex subset \( \{4,5,6,7\} \) is a minimum dominating set of \( G_B(12,2) \). Thus \( \{1,2,3,4,5,6\} \) is not a minimum dominating set. Next theorem shows a sufficient condition for the domination number of \( G_B(n,d) \) to be \( \lceil n/(d+1) \rceil \) and a method of determining the dominating set of \( G_B(n,d) \).

**Theorem 1.** If there is a vertex \( x \in V(G_B(n,d)) \) satisfying,

\[
(d-1)x - \left( \left\lceil \frac{n}{d+1} \right\rceil - l \right) \equiv 0 \pmod{n} \tag{1}
\]

for some \( l \ (0 \leq l \leq (d+1)[n/(d+1)] - n) \), then

\[
\gamma(G_B(n,d)) = \left\lceil \frac{n}{d+1} \right\rceil.
\]

**Proof.** From Lemma 1 we show the correctness of this statement by constructing a dominating set of \( G_B(n,d) \) that contains \( x \) satisfying equation (1). Let \( D = \{x, x+1, \ldots, x + \lceil n/(d+1) \rceil - 1 \} \subseteq V(G_B(n,d)) \). Then out sets of vertices in \( D \) are as follows.

\( O(x) = \{dx, dx+1, \ldots, dx + d-1\} \),

\( O(x+1) = \{d(x+1), d(x+1) + 1, \ldots, d(x+1) + d-1\} \),

\[ \vdots \]

\( O(x + \lceil n/(d+1) \rceil - 1) = \{d(x + \lceil n/(d+1) \rceil - 1), d(x + \lceil n/(d+1) \rceil - 1) + 1, \ldots, d(x + \lceil n/(d+1) \rceil - 1) + d-1\} \).
Thus \(|O(D)| = d \left\lceil \frac{n}{d+1} \right\rceil\), and \(|D| = \left\lceil \frac{n}{d+1} \right\rceil\). Since \(x\) satisfies equation (1),
\[
dx \equiv x + \left\lfloor \frac{n}{d+1} \right\rfloor - l \pmod{n}.
\]
Therefore \(|D \cap O(x)| = l\), and by the condition for \(l\), we obtain \(\left\lfloor \frac{n}{d+1} \right\rfloor - l \geq 0\).

\[
|D| + |O(D)| - |D \cap O(x)|
= \left\lfloor \frac{n}{d+1} \right\rfloor + d \left\lfloor \frac{n}{d+1} \right\rfloor - l
\geq \left\lfloor \frac{n}{d+1} \right\rfloor + d \left\lfloor \frac{n}{d+1} \right\rfloor - \{(d + 1) \left\lfloor \frac{n}{d+1} \right\rfloor - n\}
= n.
\]

Hence \(D\) is a minimum dominating set of \(G(n,d)\).

We obtain the following statement for a special case Theorem 1.

**Corollary 1.** For \(G_B(n,d), \gamma(G_B(n,d)) = 1\) if and only if \(d \geq n\) or \(n\) is odd and \(d = n - 1\).

**Proof.** We only show the necessary condition because sufficiency is obvious. Since \(\gamma(G_B(n,d)) = 1\), there is a vertex \(x \in V(G_B(n,d))\) such that \(|O[x]| \geq n\). If \(|O(x)| \geq n\) then \(d \geq n\) because of the definition of \(G_B(n,d)\). If \(|O(x)| = n - 1\) then \(d = n - 1\) and \(x \in V(G_B(n,d))\) has no loop. Hence, we obtain \(x + 1 \equiv (n - 1)x \pmod{n}\). Therefore \(2x + 1 \equiv 0 \pmod{n}\), then \(n\) is odd.

From Theorem 1, we see the following corollary for the solution of \(G_B(n,d)\).

**Corollary 2.** Let \(d + 1 | n\). Then if there is a vertex \(x \in V(G_B(n,d))\) such that
\[
(d - 1)x - \frac{n}{d+1} \equiv 0 \pmod{n},
\]
then there is a solution \(S\) of \(G_B(n,d)\) such that
\[
|S| = \frac{n}{d+1}.
\]

We prepare the reversal of a digraph \(G\), denoted by \(G^{-1}\). The reversal of a digraph \(G\) is defined by \(V(G^{-1}) = V(G)\), \((x,y) \in A(G^{-1})\) if and only if \((y,x) \in A(G)\). For generalized de Bruijn digraphs, \(G_B(n,d)\) is not necessarily isomorphic to \(G_B(n,d)^{-1}\). For de Bruijn digraph \(B(d,D)\), we can see that \(B(d,D) \cong B(d,D)^{-1}\). Then there is one to one correspondence between the dominating set and the absorbant in \(B(d,D)\). For \(B(d,D)\), the minimum cardinality of an absorbant is equal to the domination number. Therefore, we obtain the following statement.

**Corollary 3.** For \(B(d,D)\), both the minimum cardinalities of the absorbant and the dominating set are \(\left\lceil \frac{d^D}{d+1} \right\rceil\).
Proof. We first assume that $D$ is even. Then the value $l$ in Theorem 1 takes an integer between 0 to $d$ and from the definition of $B(d, D)$, $d \geq 2$. We put
$$x \equiv d^{D-2} + d^{D-4} + \cdots + d^{D-2k} + \cdots + d^2 + 1 \pmod{d^D},$$
then $x$ satisfies equation (1) in Theorem 1. Next let us assume that $D$ is odd, then the value $l$ in Theorem 1 takes an integer between 0 to $l$. We put
$$x \equiv d^{D-2} + d^{D-4} + \cdots + d^{D-2k} + \cdots + d^3 + d \pmod{d^D},$$
then $x$ satisfies equation (1) in Theorem 1. $\square$

Minimum dominating set is not necessarily a set of vertices whose labels are consecutive. For $G_B(10, 5)$, both $\{1, 2\}$ and $\{3, 6\}$ are minimum dominating sets of $G_B(10, 5)$. The set $\{1, 2\}$ is constituted of consecutive vertices and the set $\{3, 6\}$ is not constituted of consecutive vertices. Thus Theorem 1 gives a sufficient condition but not a necessary condition. It is not easy to find all minimum dominating sets of $G_B(n, d)$. We can not determine that domination number of a given $G_B(n, d)$ is equal to $\lceil n/(d+1) \rceil$ only using Theorem 1 because $G_B(n, d)$ may have a minimum dominating set which is constituted of non-consecutive vertices and whose cardinality is equal to $\lceil n/(d+1) \rceil$. Then we give a necessary and sufficient condition for $G_B(n, d)$ to have a minimum dominating set $S$ with $|S| = \lceil n/(d+1) \rceil$ and constituted of consecutive vertices. We call such $S$ a consecutive minimum dominating set.

Theorem 2. $G_B(n, d)$ satisfies one of the following four conditions if and only if there is a vertex $x \in V(G_B(n, d))$ such that $\{x, x+1, \ldots, x+\lceil n/(d+1) \rceil - 1\}$ is a consecutive minimum dominating set.

Conditions
(i) $d + 1 \nmid n$,
(ii) $d + 1 | n$ and $d$ is even,
(iii) $2(d+1) | n$ and $d$ is odd and $4 | d+1$,
(iv) $2^k(d+1) | n$ and $d$ is odd and $d+1$ contains only one $2$ as a factor.
(d $- 1 = 2^k s, s$ is odd)

Proof. It is obvious that $\gamma(G_B(n, d)) = 1$, when $d \geq n$. Suppose that $d < n$. We show that there is a vertex $x$ in $V(G_B(n, d))$ such that
$$x + \lceil n/(d+1) \rceil \equiv dx + i \pmod{n}, \quad 0 \leq i \leq \min\{d-1, \lceil n/(d+1) \rceil (d+1) - n\}.$$  \hspace{2cm} (2)

From equation (2), we obtain
$$(d-1)x \equiv \lceil n/(d+1) \rceil - i \pmod{n}. $$ \hspace{2cm} (4)

For equation (3), the value $i$ takes an integer between 0 and $d - 1$ if $d + 1$ is not a divisor of $n$. If $d + 1 | n$, then $i = 0$. $G_B(n, d)$ has a consecutive dominating set $S$
with $|S| = \lceil n/(d+1) \rceil$ if and only if there is a vertex $x \in V(G_B(n,d))$ satisfying equation (4). If $d + 1$ is not a divisor of $n$, let $(d-1,n) = s$. Then there are $q$ and $r$, $(0 \leq r < s)$ such that $\lceil n/(d+1) \rceil = qs + r$. We can put $r = i$, from the ranges of $r$ and $i$. Thus there are $x$ and $i$ satisfying equations (2) and (3) in $\mathbb{Z}_n$. Next, if $d + 1$ is a divisor of $n$,
\[(d-1)x \equiv n/(d+1) \pmod{n}\]by equation (4). The equation (5) has a solution if and only if
\[((d-1),n)|n/(d+1).\]
Hence if $d$ is even, there is an integer $x$ satisfying equation (5) in $\mathbb{Z}_n$. Whereas if $d$ is odd, $d - 1, d + 1$ and $n$ are even integers. We put $d + 1 = 2^h \alpha$ ($h \leq 2$ is an integer, $\alpha$ is an odd integer), then $d - 1 = 2(2^{h-1}\alpha - 1)$. We obtain $2|d - 1$ and $4 \mid d - 1$, since $2^{h-1}\alpha - 1$ is odd. Therefore $n$ is divided by $2^{h+1}$. Then $n$ is divided by $2(d+1)$.

For $d + 1 = 2^h \alpha$, we consider $h = 1$, then put $d - 1 = 2(\alpha - 1) = 2^k \beta$, ($\beta$ is odd). $2^{h+1}$ divides $n$ by (6). Therefore $n$ is divided by $2^h(d+1)$. Conversely if $G_B(n,d)$ has a consecutive minimum dominating set $\{x, x+1, \ldots, x+\lceil n/(d+1) \rceil - 1\}$, then $x$ satisfies equation (4). Therefore $n$ and $d$ satisfy one of above four conditions.

We consider whether there exists $G_B(n,d)$ that does not have a minimum consecutive dominating set and $\gamma(G_B(n,d)) = \lceil n/(d+1) \rceil$. We show that if $\lceil n/(d+1) \rceil = 1, 2,$ or $3$, there does not exist $G_B(n,d)$ that does not have a minimum dominating set and $\gamma(G_B(n,d)) = \lceil n/(d+1) \rceil$.

**Theorem 3.** For domination number of generalized de Bruijn digraphs, we obtain

\[
\gamma(G_B(2s,2s-1)) = 2, \quad (7)
\]
\[
\gamma(G_B(8s - 4,4s - 3)) = 3, \quad (8)
\]
and
\[
\gamma(G_B(6s,2s-1)) = 4, \quad (9)
\]
where $s$ is a natural number.

**Proof.** For equation (7), each vertex of $G_B(2s,2s-1)$ has a loop. Then $\gamma(G_B(2s,2s-1)) \neq 1$

For equation (8), let $\gamma(G_B(8s - 4,4s - 3)) = 2$ and a vertex $x$ is contained in a dominating set of $G_B(8s - 4,4s - 3)$. Then $x + 1$ or $x + 4s - 2$ is contained in the dominating set of $G_B(8s - 4,4s - 3)$. If the dominating set is $\{x, x + 1\}$, then

\[(4s - 3)x \equiv x + 2 \pmod{8s - 4},\]
\[4(s - 1)x \equiv 2 \pmod{8s - 4}.\]
Therefore we have a contradiction. If the dominating set is \( \{ x, x + 4s - 2 \} \), then we obtain

\[(4s - 3)x \equiv x + 1 \pmod{8s - 4} \tag{10} \]

or

\[(4s - 3)(x + 4s - 2) \equiv x + 1 \pmod{8s - 4}. \tag{11} \]

If equation (10) holds, then \((4s - 4)x \equiv 1 \pmod{8s - 4}\). This is a contradiction.

If equation (11) holds, then \((4s - 4)x \equiv 4(5s - 4s^2) - 5 \pmod{8s - 4}\). This is also a contradiction.

For equation (9), let us assume that \( \gamma(G_b(6s, 2s - 1)) = 3 \). We should consider three cases: (1) any two vertices are not consecutive in the dominating set, (2) exactly two vertices are consecutive in the dominating set, and (3) vertices in the dominating set are consecutive. Case (1); let a vertex \( x \) be in the dominating set, then other vertices in the dominating set are \( x + 2s \) and \( x + 4s \). Thus \( x \) satisfies

\[(2s - 1)x \equiv x + 2ts + 1 \pmod{6s} \]

for some \( t (t = 0, 1, 2) \). From this equation, we obtain

\[2(s - 1)x \equiv 2ts + 1 \pmod{6s}, \]

and this is a contradiction. Case (2); let \( x \) and \( x + 1 \) be consecutive vertices in the dominating set, then another vertex in the dominating set is \( x + 2s + 1 \) or \( x + 4s \). If the dominating set is \( \{ x, x + 1, x + 2s + 1 \} \), then

\[
\begin{cases}
(2s - 1)(x + 2s + 1) \equiv x + 2 \pmod{6s}, \\
(2s - 1)(x + 1) + 2s - 2 \equiv x - 1 \pmod{6s}.
\end{cases}
\]

By these simultaneous equations, we obtain

\[4s(s - 1) \equiv 1 \pmod{6s}. \]

Then this is a contradiction.

If the dominating set is \( \{ x, x + 1, x + 4s \} \) then

\[(2s - 1)(x + 4s) + 2s - 2 \equiv x - 1 \pmod{6s}. \]

By this equation, we obtain

\[2(s - 1)x \equiv 8s - 8s^2 + 1 \pmod{6s} \]

and this is a contradiction.

Case (3); let \( x, x + 1 \) and \( x + 2 \) constitute a dominating set. Then

\[(2s - 1)x \equiv x + 3 \pmod{6s}. \]

Hence we obtain

\[2(s - 1)x \equiv 3 \pmod{6s}, \]

and this is also a contradiction. \( \square \)
4 Domination Numbers of Generalized Kautz Digraphs

We can see the following lemma similar to Lemma \[1\].

\[\text{Lemma 3. } \gamma(G_K(n, d)) \geq \left\lceil \frac{n}{d+1} \right\rceil. \]

Besides the difficulty of determining domination numbers of generalized de Bruijn digraphs, it is easy to determine the domination number of a generalized Kautz digraph \(G_K(n, d)\).

\[\text{Theorem 4. } \gamma(G_K(n, d)) = \left\lceil \frac{n}{d+1} \right\rceil \]

\[\text{Proof. } D = \{0, 1, 2, \ldots, \left\lfloor \frac{n}{d+1} \right\rfloor - 1\} \subseteq V(G_K(n, d)) \text{ is a dominating set of } G_K(n, d). \]

5 Conclusion

We investigated domination numbers of generalized de Bruijn digraphs and generalized Kautz digraphs. Consecutive dominating set is important in our investigations. Finding a dominating set of a generalized de Bruijn digraph is more difficult than that of generalized Kautz digraph. We conjecture that there is a consecutive minimum dominating set for \(G_B(n, d)\).

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A Notion of Cross-Perfect Bipartite Graphs

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Abstract. In this note, we consider four quantities defined on a bipartite graph $B$: the cross-chromatic index $\chi^*(B)$, the biclique number $w^*(B)$, the cross-free matching number $m^*(B)$ and the biclique edge covering number $\beta^*(B)$. We mention several applications of these numbers and define a notion of cross-perfect bipartite graphs. A duality between these numbers for the class of cross-perfect graphs is examined.

Keywords: bipartite graph; perfect graph; integral polytope.

1 Definitions

Given a bipartite graph $B = (U \cup V, E)$, two non-adjacent edges $e, e' \in E$ with $e = (u_1, v_1)$ and $e' = (u_2, v_2)$ are said to form a cross if $(u_1, v_2) \in E$ and $(u_2, v_1) \in E$. Two edges are said to be cross-adjacent if either they are adjacent (i.e. share a common node) or they form a cross. A cross-free matching in $B$ is a set of edges $E' \subseteq E$ with the property that no two edges $e, e' \in E'$ are cross-adjacent. A cross-free coloring of $B$ is a coloring of the edge set $E$ subject to the restriction that no pair of cross-adjacent edges has the same color. A set of edges $E' \subseteq E$ is said to form a biclique if the induced subgraph corresponding to $E'$ is a complete bipartite subgraph of $B$. A biclique edge cover for $B$ is a covering of the edge set $E$ by bicliques.

We now consider the following four quantities on $B$: (i) The cross-chromatic index, $\chi^*(B)$, of $B$ which is the minimum number of colors required to get a cross-free coloring of $B$ (ii) The biclique number, $w^*(B)$, of $B$ defined as the cardinality of the maximum edge biclique in $B$ (iii) The cross-free matching number of $B$, $m^*(B)$, defined as the edge cardinality of the maximum cross-free matching in $B$ and (iv) The biclique edge covering number of $B$, $\beta^*(B)$, defined as the minimum number of bicliques required to cover the edges of $B$.

This note is concerned with a duality between the four numbers defined above. The duality, as stated here, resembles the classical duality between the four quantities, namely, stability number $\alpha(G)$, clique partition number $\theta(G)$, chromatic number $\gamma(G)$ and the maximum clique number $\omega(G)$ for a perfect graph $G$ [1][2][9]. In fact, we show that the duality between $w^*(B)$, $\chi^*(B)$, $m^*(B)$ and $\beta^*(B)$ can be studied by investigating when a certain class of graphs is perfect.

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2 Applications and Complexity Results

The problem of finding a maximum cross-free matching in a bipartite graph is discussed in [16]. This problem was studied as a special case of a more general problem of finding alternate cycle free matchings in bipartite graphs. Alternate cycle free matchings were first investigated in [10] in the context of dependence graphs of matroids: if $\mathcal{M}$ is a matroid with basis $B$, finding a maximum cardinality alternating cycle free matching is related to finding a basis of $\mathcal{M}$ for which the intersection with $B$ is of minimum cardinality. Another application studied in [3] and mentioned in [16] concerns the computing of the setup number or step number $\nu(G)$ of an acyclic directed graph $G$.

The problem of finding a maximum edge cardinality biclique has applications in manufacturing in the computer industry. This application is discussed in detail in [5]. In short, the relationship between a set of products and a set of components which constitute the products can be represented by a bipartite graph. One way to reduce manufacturing lead times is to reduce the final assembly times for the products by creating sub-assemblies. Finding good sub-assemblies in this context can be done by finding large edge-cardinality bicliques in the bipartite graph. Another interesting application, also described in [5], occurs in the area of formal concept analysis [7].

Covering the edges of a bipartite graph by bicliques has an interesting application [4] is combinatorial chemistry: Split Synthesis [6][11], is a method of choice to build libraries for combinatorial chemistry. The basis for this technology is that molecules can be grown with one end tethered to a resin bead, while the other end is extended one residue at a time in parallel across a set of beads. Split synthesis can be modeled by a directed acyclic graph in which each node represents a grow step on a particular subset of beads. In [4], it is shown that optimizing split synthesis construction is intimately related to covering the edges of a bipartite graph by bicliques.

Finding $\beta^*(B)$ is NP-hard [8]. In [16], it is shown that finding $m^*(B)$ is NP-hard. Recently, R. Peeters [15] showed that finding $w^*(B)$ is also NP-hard. To the best of our knowledge, the complexity of finding $\chi^*(B)$ is open.

3 Formulations and Results

Clearly, a biclique cannot have more than one edge from a cross-free matching. This immediately gives the following integer programming formulation for $w^*(B)$

$$w^*(B) = \max\{1x : K^*x \leq 1, x \in \{0, 1\}^n\}$$

where $n = |E|$ and $K^*$ is a $(0, 1)$ incidence matrix whose rows correspond to maximal cross-free matchings of $B$ and whose columns correspond to the edges of $B$. It is also easy to see that

$$\chi^*(B) = \min\{1y : yK^* \geq 1, y \in \{0, 1\}^m\}$$
The fractional biclique polytope of $B$ is $P = \{ x \in \mathbb{R}^n_+ : K^* x \leq 1 \}$. We denote the convex hull of feasible integer solutions of $P$ by $P_I$. We generalize the definitions of $w^*$ and $\chi^*$ to consider binary objective functions. For $c \in \{0, 1\}^n$

$$w^*(B_c) = \max\{cx : K^* x \leq 1, x \in \{0, 1\}^n\}$$

$$\chi^*(B_c) = \min\{1y : yK^* \geq c, y \in \{0, 1\}^m\}$$

Note that the definitions of $w^*(B)$ and $w^*(B_c)$ differ only in the objective function. The constraint matrix $K^*$ is the same for both. Also, $w^*(B_c)$ does not correspond to the optimal edge cardinality biclique in the subgraph (either edge-induced or node-induced) corresponding to the vector $c$. This is an important distinction between the definition of $w^*(B_c)$ and the corresponding definition for the maximum clique in a graph: For a graph $G$, the formulation for the maximum clique, $\alpha(G)$, generalized to binary objective functions, can be written as

$$\alpha(G_c) = \max\{cx : K x \leq 1, x \in \{0, 1\}^p\}$$

where $p = |V(G)|$ and $K$ is the clique matrix of $G$ \cite{14}. Here, $\alpha(G_c)$ corresponds to the maximum clique on the node-induced subgraph of $G$ corresponding to $c$. A similar observation can be made about $\chi^*(B_c)$.

**Definition:** $B$ is cross-perfect if $w^*(B_c) = \chi^*(B_c)$ for all vectors $c \in \{0, 1\}^n$.

**Theorem 31** $B$ is cross-perfect iff $P$ is integral.

**Proof.** Consider the line graph, $L(B)$, of $B$. For every 4-hole in $L(B)$, add the two cords. Let the new graph be called $L'(B)$. We refer to $L'(B)$ as the modified line graph of $B$. Then, $E' \subseteq E$ is a cross-free matching in $B$ iff $V(E')$ is an independent set in $L'(B)$. The clique polytope for $L'$ is $Q = \{ x : Sx \leq 1, x \in \mathbb{R}^n_+ \}$ where the rows of $S$ correspond to maximal independent sets and the columns correspond to the nodes of $L'$. But $Q = P$. Thus, $B$ is cross-perfect implies that $w(H) = \gamma(H)$ for every node induced subgraph $H$ of $L'$ where $w(H)$ denotes the cardinality of the maximum clique in $H$ and $\gamma(H)$ is the chromatic number of $H$. Then, $L'$ is perfect by the weak perfect graph theorem \cite{12, 13}. It follows that $Q$ and hence $P$ is integral. Conversely, if $P$ is integral, so is $Q$ and hence $w(H) = \gamma(H)$ for every node induced subgraph $H$ of $L'$. Thus, $w^*(B_c) = \chi^*(B_c)$ for all vectors $c \in \{0, 1\}^n$ and $B$ is cross-perfect.

**Note 1:** An alternative proof of Theorem 3.1 was suggested by Prof. R. Chandrasekaran: The column-intersection graph of the matrix $K^*$ is the complement graph of the modified line graph $L'(G)$ defined above. The definition of the cross-perfectness and the perfect graph theorem \cite{12} then imply that $P$ is integral.

**Note 2:** Even though the proof of Theorem 3.1 uses the correspondence between the biclique number $w^*$ and the clique number $w$, the structure of these two problems is fundamentally different. In fact, even the complexity question of computing $w^*$ was open until the recent result of Peeters \cite{15}. Theorem 3.1 and Theorem 3.4 (given below) are thus remarkable since they mention a duality,
similar to the classical duality in the perfect graph literature, for fundamentally different problems.

In general, the biclique polytope, \( P \), is not integral. This follows because if \( P \) were integral, so would the clique polytope, \( Q \), of the graph \( L'(B) \) which in turn would mean that \( L'(B) \) is perfect. However, it is easy to show a \( L'(B) \) which contains an odd hole of length 5 as a node-induced subgraph (see figure below).

**Fig. 1.** A bipartite graph whose biclique polytope is not integral

**Corollary 32** \( B \) is cross-perfect iff the modified line graph \( L'(B) \) is perfect.

This observation suggests that the property of cross-perfectness in bipartite graphs can be studied by studying perfectness in modified line graphs. If \( B \) is cross-perfect then, by Theorem 3.1, the constraints \( K^*x \leq 1 \) are sufficient to describe \( P_I \). Note that every clique in \( L' \) does not correspond to a biclique in \( B \). However, every maximal clique in \( L' \) does correspond to a biclique. Hence, a maximum clique, which is also a maximal clique, corresponds to a maximum biclique in \( B \). This observation immediately gives us the following result.

**Theorem 33** If \( B \) is cross-perfect then \( w^*(B) \) and \( \chi^*(B) \) are polynomially solvable.

Every independent set in \( L' \) corresponds to a cross-free matching in \( B \). Consider a covering, \( C \), of the nodes of \( L' \) by cliques. A clique \( c^* \) in \( C \) either corresponds to a biclique in \( B \) or if not, \( c^* \) is contained in a maximal clique which corresponds to a biclique in \( B \). Thus, \( \alpha(L') = m^*(B) \) and \( \theta(L') = \beta^*(B) \). Let \( H \) be a node induced subgraph of \( L' \) and \( c \in \{0,1\}^n \) be the incidence vector of the nodes corresponding to \( H \). Then, it is easy to see that \( \alpha(H) = m^*(B_c) \) and \( \theta(H) = \beta^*(B_c) \). This correspondence along with the weak perfect graph theorem \[12\] [13] gives us the following result.

**Theorem 34** The following are equivalent
1. \( w^*(B_c) = \chi^*(B_c) \) for all vectors \( c \in \{0,1\}^n \).
2. \( m^*(B_c) = \beta^*(B_c) \) for all vectors \( c \in \{0,1\}^n \).

**Theorem 35** Node-induced subgraphs of cross-perfect graphs are cross-perfect
Proof. The fractional biclique polytope, $P_H$, corresponding to a node induced subgraph $H$ of $B$ is a face of $P$. Hence if $P$ is integral, so is $P_H$.

**Corollary 36** If $B$ is cross-perfect then $w^*(H) = \chi^*(H)$ for all node induced subgraphs $H$ of $B$.

Getting an efficient recognition algorithm for cross-perfect bipartite graphs and the validity of the strong perfect graph conjecture [2][9] for modified line graphs (defined in Theorem 3.1) are outstanding open questions. Also, it is not known whether the class of modified line graphs of cross-perfect bipartite graphs is a new class of perfect graphs.

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Some Results on Orthogonal Factorizations*

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Abstract. Consider a graph $G = (V,E)$ with an $[a,b]$–factorization $F = \{F_1,F_2,...,F_m\}$. It is proved in this paper that:

1. there is an $m$–matching of $G$ to which $F$ is orthogonal if $n = |V(G)| \geq (2 + \frac{b}{a})(m - 1)$ for $b \geq 2a$ and $n \geq 3.26m$ for $b = a$;
2. if $\sqrt{2b} \leq a \leq b$, then for any given edge $e$ of $G$, there is a $[1,a]$–subgraph $H$ of $G$ such that $e$ is included in $H$ and $F$ is orthogonal to $H$.

1 Introduction

All graphs considered are undirected graphs without loops and without multiple edges. Let $G$ be a graph. we denote by $V(G)$ and $E(G)$ the vertex set and edge set of $G$ respectively. Denote by $d_G(v)$ the degree of $v$ in $G$. Let $a$ and $b$ be two integers. An $[a,b]$–factor of $G$ is a spanning subgraph $F$ of $G$ such that $a \leq d_F(v) \leq b$ for every $v \in V(F)$. In particular, if $a = b = k$, then an $[a,b]$–factor is a $k$–factor. For simplicity and in case when there is no ambiguity, we denote a subgraph and its edge set with the same symbol. An $[a,b]$–factorization $F = \{F_1,F_2,...,F_m\}$ of $G$ is a partition of $E(G)$ into edge disjoint $[a,b]$–factors $F_1,F_2,...,F_m$. Similarly, we define $k$–factorization as a partition of $E(G)$ into $k$–factors. Let $H$ be a subgraph of $G$ with $m$ edges and $F = \{F_1,F_2,...,F_m\}$ an $[a,b]$–factorization of $G$, then $F$ is called orthogonal to $H$ if each $F_i, 1 \leq i \leq m$ has exactly one edge in common with $H$. Other definitions and terminologies can be found in [2].

Alspach B. posed the following conjecture:

Conjecture 1. [1] Let $G$ be a $2d$–regular graph, $F = \{F_1,F_2,...,F_m\}$ be a $2$–factorization of $G$, then there exists a $d$–matching $M$ of $G$ to which $F$ is orthogonal.

G.Z.Liu et al. [4] proved that the conjecture is true when $|V(G)| \geq 4d - 5$ and extend the question to $k$–factorizations and $[1,k]$–subgraphs. Later, M. Kouider et al. [3] improved the result of [4] to $|V(G)| \geq 3.23d$.

In this paper, we improve the above results and extend them to $[a,b]$–factorization. The main techniques of this paper come from those in [4] and [3].

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Let $A$ and $B$ be two disjoint subsets of $V(G)$, then for any $1 \leq i \leq m$, denote by $E_i(A,B)$ the set of edges of factor $F_i$ with one end in $A$ and the other in $B$, by $E_i(A)$ the set of edges of factor $F_i$ with both ends in $A$. Let $k_1, ..., k_j$ be $j$ different indexes between 1 and $m$, then denote by $E_{k_1, ..., k_j}(A,B)$ the set of edges of all the factors $F_{k_1}, ..., F_{k_j}$ with one end in $A$ and the other in $B$.

2 Main Theorems and Proofs

**Theorem 1.** Let $a, b$ and $m$ be positive integers with $a \leq b$. Let $G$ be a graph with an $[a,b]$–factorization $F = \{F_1, F_2, ..., F_m\}$. Then there is an $m$–matching of $G$ to which $F$ is orthogonal if $n = |V(G)| \geq (2 + \frac{b}{a})(m - 1)$ for $b \geq 2a$ and $n \geq 3.26m$ for $b = a$.

*Proof.* We apply induction on $m$. When $m = 1$, the theorem is obviously true. Suppose the theorem is true for $m - 1$. We will prove that the theorem holds for $m \geq 2$. By induction hypothesis, there is an $(m-1)$–matching $M'$ of $G$ to which $F' = \{F_1, F_2, ..., F_{m-1}\}$ is orthogonal. Assume that there is no $m$–matching of $G$ as described in theorem 1. We will see a contradiction.

W.l.g., suppose $M' = \{e_1, e_2, ..., e_{m-1}\}$, where $e_i = x_iy_i \in F_i, 1 \leq i \leq m - 1$. Let $A = \{x_i, y_i : 1 \leq i \leq m - 1\}$, $B = V(G) \setminus A$. Then $|A| = 2(m - 1)$, $|B| = n - 2(m - 1)$. Since there is no $m$–matching of $G$ satisfying the conclusion of theorem 1 according to our assumption, we have that $E_m(B) = \phi$. Then $E_m = E_m(A,B) + E_m(A)$.

We now reindex the edges in $M'$ and let $s$ denote the maximum index subject to the following two conditions,

1. there exist two distinct vertices $x'_1$ and $y'_1$ in $B$ satisfying that $x_1x'_1$ and $y_1y'_1$ are edges in $F_m$;
2. for each index $i$ between 2 and $s$, there exist two distinct vertices $x'_i$ and $y'_i$ satisfying that $x_ix'_i$ and $y_iy'_i$ are edges in $F_{i-1}$. All the $\{x'_i, y'_i\}(i \geq 1)$ are disjoint.

Consider the subset $S = \{e_1, ..., e_s\}$ of $M'$. We claim that $S$ is nonempty. For there are always some edges, say $e_j = x_jy_j$ in $M'$ which are adjacent to two independent edges in $E_m(A,B)$. Otherwise, for each edge $e_j = x_jy_j \in M'$, either the edges in $F_m(A,B)$ are only incident with one end point of $e_j$ or both the edges in $F_m(A,B)$ incident with $x_j$ and $y_j$ end in the same vertex in $B$. Therefore, we always have $a|B| \leq b|M'| = b|A|$, and subsequently we have

$$n = |A| + |B| \leq (1 + \frac{b}{2a})|A| = (2 + \frac{b}{a})(m - 1),$$

a contradiction to our assumption.

For each index $k$ between 1 and $s$, denote by $A_k$ the set of the ends of $e_1, e_2, ..., e_k$, by $A'_k$ the complement of $A_k$ in $A$; denote by $B_k$ the set of $x'_i, y'_i, 1 \leq i \leq k$, by $B'_k$ the complement of $B_k$ in $B$. Then we have:

$$|A_k| = |B_k| = 2k, |A'_k| = 2(m - 1 - k), |B'_k| = n - 2m + 2 - 2k.$$  

(2)
Since according to our assumption there is no matching of $G$ to which $F$ is orthogonal, we have

$$E_S(A_S, B'_S) = E_S(B_S, B'_S) = E_S(B'_S, B'_S) = \emptyset.$$  \hfill (3)

Besides, due to the maximality of $s$, similar as the discussion before, any edge $x$ with $x \in E_{M'}(A'_S)$ is adjacent to at most one independent edge in $E_S(A_S, B'_S)$. Therefore we have:

$$|B'_S| \leq \frac{b}{2a} |A'_S|,$$  \hfill (4)

which implies that:

$$n = |B'_S| + 2m - 2 + 2s \leq \frac{b}{2a} |A'_S| + 2m - 2 + 2s = (2 + \frac{b}{a})(m - 1) + (2 - \frac{b}{a})s.$$  \hfill (5)

Similarly, we have:

$$E_{k,k+1,\ldots,s}(A_k, B'_k) = \emptyset,$$
$$E_{k,k+1,\ldots,s}(B_k, B'_k) = \emptyset,$$
$$E_{k,k+1,\ldots,s}(B'_k, B'_k) \subseteq \{x'_iy'_i| k + 1 \leq i \leq s\},$$
$$E_{k,k+1,\ldots,s}(A_k \cup B_k, A'_k \cup B'_k),$$
$$\subseteq \{x'_iy'_i, y'_i, x'_i| 1 \leq i \leq k\} \cup \{e_k\},$$  \hfill (6)

Subsequently we have:

$$|E_{k,k+1,\ldots,s}(A'_k, A'_k)| \geq a|A_k|(s - k + 1) - |A_k| - 2,$$
$$|E_{k,k+1,\ldots,s}(B'_k, B'_k)| \geq a|B_k|(s - k + 1) - 2|B_k|,$$
$$|E_{k,k+1,\ldots,s}(B'_k, A'_k)| \geq a|B'_k|(s - k + 1) - 2(s - k).$$  \hfill (7)

And therefore,

$$|E_{k,k+1,\ldots,s}(A'_k, A_k \cup B)| \geq a(s - k + 1)(|A_k| + |B|)$$
$$- (6k + 2) - 2(s - k).$$  \hfill (8)

On the other hand, counting the edges in $E_{k,k+1,\ldots,s}(A'_k, A_k \cup B)$, we have that:

$$|E_{k,k+1,\ldots,s}(A'_k, A_k \cup B)| \leq b|A'_k|(s - k + 1) - 2(s - k).$$  \hfill (9)

Subsequently we have,

$$b|A'_k|(s - k + 1) - 2(s - k) \geq a(s - k + 1)(|A_k| + |B|)$$
$$- (6k + 2) - 2(s - k).$$  \hfill (10)

Therefore,

$$a(s - k + 1)(2k + n - 2m + 2) - b(s - k + 1)(2m - 2k) \leq 6k + 2.$$  \hfill (11)

Hence, we have:

$$n \leq \frac{2m(a + b)}{a} - \frac{2(a + b)}{a} - \frac{2k(a + b)}{a} + \frac{6k + 2}{a(s - k + 1)}.$$  \hfill (12)

Now we have two cases to discuss.
Case 1. $b \geq 2a$. 

From (5) we have that

$$n \leq (2 + \frac{b}{a})(m - 1), \quad (13)$$

a contradiction to our assumption.

Case 2. $b = a$.

From (5), we have

$$n \leq 3m - 3s. \quad (15)$$

From (5), we have $s \geq n - 3m + 3$ and subsequently we have

$$5n \leq (10 + \frac{10b}{a})m - (10 + \frac{10b}{a} - \frac{6}{a})s + 30 - \frac{22}{a} + \frac{30b}{a}. \quad (16)$$

Therefore,

$$n \leq \frac{40a + 40b - 18}{15a + 10b - 6} m - \frac{4}{a} \leq \frac{80a - 18}{25a - 6} m \leq 3.26m, \quad (17)$$

a contradiction to our assumption that $n \geq 3.26m$.

Thus completes the proof of Theorem 1. \qed

**Corollary 1.** Let $G$ be a $kd$–graph with vertex number $n \geq 3.26d$ and let $F = \{F_1, F_2, ..., F_d\}$ be a $k$–factorization of $G$. Then $G$ has a $d$–matching to which $F$ is orthogonal.

*Proof.* We can prove this corollary by setting $a = b = k$ in Theorem 1. \qed

Corollary 1 greatly improved G.Z.Liu et al.’s results concerning orthogonal $k$–factorizations in [4].

**Theorem 2.** Let $a, b$ and $l$ be integers such that $\sqrt{2b} \leq a \leq b$. Let $G$ be a graph, $F = \{F_1, F_2, ..., F_l\}$ be an $[a, b]$–factorization of $G$. Then for every given edge $e$ of $G$, $G$ has a $[1, a]$–subgraph $H$ such that $e \in H$ and to $H F$ is orthogonal.
Proof. We use induction on $l$. When $l = 1$, the theorem is trivial. Suppose the theorem is true for $l \leq m - 1$, we will prove it is true for $l = m$.

W.l.g., suppose $e \notin F_m$. By induction assumption, $G$ has a $[1,a]$–subgraph $H'$, such that $e \in H'$ and that to $H'$, $\mathcal{F}' = \{F_1, F_2, \ldots, F_{m-1}\}$ is orthogonal. Suppose $G$ has no satisfying $[1,a]$–subgraphs, we will see a contradiction to our given conditions of theorem 2.

Let

$$S = \{x_i, y_i : x_i y_i \in E(H')\}, T = V(G) \setminus S.$$  \hfill (18)

Let

$$S_l = \{x \in S \text{ and } d_{H'}(x) < a\},$$

$$S_a = \{x \in S \text{ and } d_{H'}(x) = a\}.$$  \hfill (19)

Since $G$ has no satisfying $[1,a]$–subgraphs, we have that $E_m = E_m(S_a, V(G))$. On the other hand, $E_m(T, S_l) = \emptyset, E_m(T, T) = \emptyset$. Therefore,

$$E_m(T, V(G)) = E_m(T, S_a),$$  \hfill (20)

and subsequently,

$$|E_m(T, V(G))| \leq |E_m| = |E_m(S_a, V(G))|. $$  \hfill (21)

Notice that

$$|E_m(T, V(G))| \geq a|T|, |E_m(S_a, V(G))| \leq b|S_a|,$$  \hfill (22)

we have

$$a|T| \leq |E_m(T, V(G))| \leq |E_m(S_a, V(G))| \leq b|S_a|,$$  \hfill (23)

and subsequently we have

$$|T| \leq \frac{b}{a}|S_a|.$$  \hfill (24)

Now we have two cases to discuss:

Case 3. $a = 2$.

In this case, $H'$ is a $[1,2]$–subgraph of $G$. Hence,

$$2(m - 1) = |S_l| + 2(|S| - |S_l|).$$  \hfill (25)

Then we have

$$|S_l| = 2(|S| - m) + 2, |S_a| = |S| - |S_l| = 2m - |S| - 2.$$  \hfill (26)

Notice that

$$|T| = n - |S| \leq \frac{b}{a}|S_a| = \frac{b}{a}(2m - |S| - 2),$$  \hfill (27)

we have

$$n \leq \frac{2b}{a}m - \frac{2b}{a} - \frac{b}{a}|S| < \frac{2b}{a}m.$$  \hfill (28)

On the other hand, since $n \geq ma + 1$, we have

$$ma + 1 < \frac{2b}{a}m,$$  \hfill (29)

which implies $a < \sqrt{2b}$, a contradiction to $a \geq \sqrt{2b}$. 

Case 4. $a \geq 3$.

Since $H'$ has $m - 1$ edges, we have $a|S_a| \leq 2(m - 1)$ and therefore $|S_a| \leq \frac{2(m - 1)}{a}$. Recall that $|T| \leq \frac{b}{a}|S_a|$, $|T| = n - |S|$, we have

$$n \leq |S| + \frac{b}{a}|S_a| \leq 2m - 2 + \frac{2b(m - 1)}{a^2} = (2 + \frac{2b}{a^2})m - 2 - \frac{2b}{a^2}. \quad (30)$$

Since $a \geq \sqrt{2b}$, we have $\frac{2b}{a^2} \leq 1$ and therefore $n \leq 3m - 2 = ma + 1$, a contradiction to $n \geq ma + 1$.

Theorem 2 is proved. \(\square\)

Remark 1. We can derive the results concerning $[1, k]$–subgraph in [4] by setting $a = b = k$ in Theorem 2.

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Cluttered Orderings for the Complete Graph

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Abstract. In a systematic erasure code for the correction of two simultaneous erasures, each information symbol must have two associated parity symbols. When implemented in a redundant array of independent disks (RAID), performance requirements on the update penalty necessitate that each information symbol be associated with no more parity symbols than the two required. This leads to a simple graph model of the erasure codes, with parity symbols as vertices and information symbols as edges. Ordering the edges so that no more than \( f \) check disks (vertices) appear among any set of \( d \) consecutive edges is found to optimize access performance of the disk array when \( d \) is maximized. These cluttered orderings are examined for the complete graph \( K_n \). The maximum number \( d \) of edges is determined precisely when \( f \leq 5 \) and when \( f = n - 1 \), and bounds are derived in the remaining cases.

1 Background

Let \( G = (V, E) \) be a graph on \( n = |V| \) vertices and \( m = |E| \) edges. Let \( d \leq m \) be a positive integer, called the window. Let \( E = \{e_0, e_1, \ldots, e_{m-1}\} \). An edge ordering of \( G \) is a permutation \( \pi \) of the edge indices \( \{0, 1, \ldots, m-1\} \). For the graph \( G \) with edge ordering \( \pi \), and window \( d \), define the \( m+1-d \) graphs \( \{G_{i}^{\pi,d} : 0 \leq i \leq m-d\} \) by setting \( G_{i}^{\pi,d} \) to be the graph containing edges \( \{e_{\pi(i)}, e_{\pi(i+1)}, \ldots, e_{\pi(i+d-1)}\} \). Then, when sets of \( d \) edges that are consecutive under the ordering \( \pi \) are accessed, the graphs \( \{G_{i}^{\pi,d}\} \) represent the possible subgraphs accessed.

In the application to be described, the cost of accessing a subgraph of \( d \) consecutive edges is measured by the sum of the number of edges and the number of vertices of nonzero degree in the subgraph. Since each has \( d \) edges, any reduction in access cost results from varying the numbers of vertices. Hence we define \( n_{i}^{\pi,d} \) to be the number of vertices of nonzero degree in \( G_{i}^{\pi,d} \). The (average) \( d \)-access cost of graph \( G \) under ordering \( \pi \) is defined to be \( \frac{\sum_{i=0}^{m-d} n_{i}^{\pi,d}}{(m+1-d)} \). The \( d \)-maximum access cost of graph \( G \) under ordering \( \pi \) is defined to be \( \max_{i=0}^{m-d} n_{i}^{\pi,d} \). Maximum access cost and average access cost measure different aspects of performance,
but evidently address similar goals. An ordering is \((d, f)\)-cluttered if it has \(d\)-maximum access cost equal to \(f\). Let \(m_d(G)\) denote the minimum value of \(f\), over all edge orderings \(\pi\) of \(G\), of the \(d\)-maximum access cost. It may happen that \(m_d(G) = m_{d+1}(G)\) (see [7], for example). We write \(c_f(G)\) for the maximum value of \(d\) for which a \((d, f)\)-cluttered ordering of \(G\) exists. When \(G = K_n\), we use the notation \(m_d(n)\) and \(c_f(n)\).

In this paper, we develop two combinatorial constructions for cluttered orderings of the complete graph \(K_n\). The first is for small windows (§3) and the second for large windows (§4). Each is limited to specific parameter selections for \(d\), \(f\), and \(n\), and hence may not provide the required solution for parameters arising in a particular application. However, we derive a powerful extension of the combinatorial labelling technique for small windows which dramatically increases its application. In section 5 we also develop an inflation technique that permits us to obtain orderings for specific parameter situations from solutions for smaller cases.

2 An Application to RAID

Secondary storage devices play an important role in system performance. Improvements in magnetic disk speed have not kept pace with improvements in CPU performance. Current systems therefore have the potential to be I/O bound. This, coupled with the requirement for high reliability and availability, has resulted in use of redundant arrays of independent disks (RAID); see [3,9,11,12,13,14]. RAID maps larger disk reads or writes to multiple smaller reads and writes. It splits (stripes) these across several disks in parallel along a parity stripe [16]. Redundant information is maintained in order to reconstruct the array in the case of disk failure. Parity computations are employed to retain the redundant information.

A catastrophic failure renders a disk unreadable; this is an erasure. A code that corrects for \(t\) erasures is a \(t\)-erasure code. One important metric in erasure codes for disk array architecture is the update penalty. This is the number of disks whose contents must be changed each time a write occurs on a single disk in the array. In a \(t\)-erasure code, the minimal update penalty is \(t\).

In the schemes that we examine, there are \(c + k\) disks. Within each stripe across these disks, \(k\) positions are reserved for information while \(c\) express parity functions of the information. The code used is systematic, because information is expressed in an unencoded format. The equations for an erasure code using parity for redundancy in disk arrays can be represented as a a parity check matrix, \(H = [P|I]\). This is a \(c \times (k + c)\) matrix where the columns of \(P\) are indexed by the information disks, the columns of \(I\) are indexed by check disks and the rows of \(P\) and \(I\) by the check disks. \(I\) is the \(c \times c\) identity matrix. A disk failure is recoverable if and only if the failed columns are linearly independent modulo 2 [10].

In this setting, for each write to an information disk, we must also update each of its check disks. Writes are expensive operations in a disk array. Update penalty
is a useful metric, but it fails to describe the observed write penalty. Consider writing information to the array. One method is to write to each information disk to be updated; then each of the information disks in their equations is read, in order to calculate the new parity; and lastly the changed parity is written to each of the check disks. This is unnecessarily expensive when writing to a small fraction of disks in an array. For large disk arrays with multiple erasure protection, when the number of disks to be written is relatively small compared with the number of information disks, all writes are implemented as a read-modify-write (see [6, 11]). To write to a disk, we only need to calculate the change to its check disks, before updating them. Therefore we read all of the check and information disks involved. Then the change to the parity is computed, the check disks are written and the new data is written. For a write involving only one information disk in a 2-erasure code, this requires six I/O accesses. The actual penalty for an update penalty of \( t \) is \( 2t + n \) accesses where \( n \) is the number of information disks. For small writes, this is less expensive than reading the entire stripe.

Our previous work on 3-erasures shows that the ordering of the parity check matrix may reduce the observed update penalty when writing across three disks at a time [5, 6]. Since information disks share check disks, we can improve the observed penalty if we impose an ordering and assume some intermediate buffering or parity logging (see [5, 17]). If data is striped sequentially along disks in a parity check matrix and these disks share check disks, the actual number of reads and writes can be reduced.

The connection to cluttered edge orderings is as follows. Columns of the parity check matrix contain two ‘1’ entries, and hence if we interpret check disks as vertices, every information disk is associated with an unordered pair of check disks, i.e. with an edge. The parity check matrix therefore represents a graph; indeed if it does correct two erasures, the graph has no multiple edges. Therefore the maximum number of information disks that can be employed corresponds to the selection of the complete graph to form the parity check matrix. If we label the first \( h \) columns of the parity check matrix with \( h \) distinct pairs on \( c \) vertices and then label the identity portion of the matrix with \( h + 1 \) through \( h + c \) we form a parity check matrix. If we then order the \( h \) columns using a cluttered ordering of the corresponding graph, and stripe data sequentially, this minimizes the maximum number of check disks affected by a write operation.

3 Wrapped Labellings

In this section, we develop a technique for producing specific \((d, f)\)-cluttered orderings of \( K_n \) when \( f \) is relatively small. We define a graph \( G(h, w) \) on \( h(w+1) \) vertices. The vertex set is \( V = \{v_{ij} : 0 \leq i < h, 0 \leq j \leq w\} \). The edge set is \( E = \{v_{ij}, v_{kj} : 0 \leq i < k < h, 0 \leq j < w\} \cup \{v_{ij}, v_{k,j+1} : 0 \leq k \leq i < h, 0 \leq j < w\} \). There are \( w \cdot h^2 \) edges; indeed, \( G(h, w) \) has a natural edge-partition into \( w \) copies of \( G(h, 1) \). A \( \rho \)-labelling of a graph \( G = (V, E) \) is a mapping \( \rho : V \rightarrow \mathbb{Z}_{2|E|+1} \) for which \( \{\pm(\rho(r) - \rho(s)) : \{r, s\} \in E\} \) contains all nonzero
Cluttered Orderings for the Complete Graph

numbers in $\mathbb{Z}_{2|E|+1}$; see [15]. These $\rho$-labellings have been extensively studied as methods of partitioning (or decomposing) the edge set of the complete graph $K_{2|E|+1}$ into copies of $G$ [11,15]. Under suitable conditions, they can also be used to constrain the graphs induced among all sets of $|E|$ consecutive edges; a simple example is given in [7]. We pursue a more general formulation next.

3.1 Wrapped Labelling of $G(h, w)$

A $\rho$-labelling of $G(h, w)$ is a wrapped $\rho$-labelling if it satisfies \{\(\alpha + \rho(v_{i0}) : 0 \leq i < h\}\} = \{\rho(v_{iw}) : 0 \leq i < h\}$ for some $\alpha$ which is relatively prime to $2|E|+1$. Figure 1 shows a wrapped $\rho$-labelling for $G(3, 2)$; here $\alpha = -2$.

**Proposition 1.** If $G(h, w)$ has a wrapped $\rho$-labelling, then there is an ordering of the edges of $K_{2wh^2+1}$ in which every set of $h^2$ consecutive edges induces a subgraph on at most $2h$ vertices, i.e. a $(h^2, 2h)$-cluttered ordering. Thus $m_{h^2}(2wh^2 + 1) \leq 2h$ and $c_{2h}(2wh^2 + 1) \geq h^2$.

**Proof.** Let $\alpha + \rho$ be the labelling obtained by adding $\alpha$ to each label assigned by $\rho$, reducing modulo $2wh^2 + 1$. Starting with $(G_{hw}, \rho)$, we can identify the vertices $\{v_{iw}\}$ with the vertices $\{v_{i0}\}$ of $(G_{hw}, \alpha + \rho)$ so that identified vertices have the same label (since the labelling is a wrapped $\rho$-labelling). Repeating this addition of $\alpha$, we encounter every translate of the labelling $\rho$ since $\alpha$ and $2wh^2+1$ are relatively prime. Indeed, in the process, every edge of $K_{2wh^2+1}$ arises in exactly one of the translated graphs constructed, since every edge difference arises exactly once in $\rho$. An example, adjoining the first four translates of $G(3, 1)$ with wrapped $\rho$-labellings $\rho, 1 + \rho, 2 + \rho,$ and $3 + \rho$, is shown in Figure 2.

Our task is now to order the edges of $G(h, 1)$. We first list the \(\binom{h}{2}\) edges on $\{v_{i0}\}$ in any order. Then we start with $\{v_{00}, v_{01}\}; \{v_{10}, v_{01}\}, \{v_{10}, v_{11}\};$ and in general, $\{v_{i0}, v_{01}\}, \ldots, \{v_{i0}, v_{11}\}$. In other words, we list the edges of the form $\{v_{i0}, v_{j1}\}$ so that $\{v_{i0}, v_{j1}\}$ precedes $\{v_{k0}, v_{l1}\}$ if and only if $i < k$, or $i = k$ and $j < \ell$. If, within every copy of $G(h, 1)$, edges are ordered in this manner, then we claim that every set of $h^2$ consecutive edges induces at most $2h$ vertices. Consider sliding a window of size $h^2$ through the first and second copies of $G(h, 1)$. Initially there are $2h$ vertices. The first $\binom{h}{2}$ transitions move the edges on $\{v_{i0}\}$ to the corresponding edges on $\{v_{i1}\}$; no new vertices are encountered.

![Fig. 1. Wrapped $\rho$-labelling of $G(3, 2)$](image-url)
Fig. 2. First four translates of $G(3, 1)$

Fig. 3. First nine windows for $G(3, 2)$

The next transition drops the only edge still incident at $v_{00}$ and adds an edge to $v_{02}$. The next edge introduced involves $v_{02}$, so no new vertex is seen. Then the last edge incident at $v_{10}$ is dropped, and an edge to $v_{12}$ added. Carrying this on, whenever a new vertex is to be introduced, say $v_{i2}$, the last remaining edge incident with $v_{i0}$ is being dropped. Thus after $h^2$ transitions, the window contains the second copy of $G(h, 1)$ in the $G(h, w)$. See Figure 3 for an illustration of the first nine windows (arranged horizontally) of the ordering of $G(3, 2)$. Transitions into all copies of $G(h, 1)$ are effected in this way. When the first $G(h, w)$ is fully seen, since the second $G(h, w)$ is attached to it as described above, we can simply continue with the same argument. Thus every set of $h^2$ consecutive edges induces at most $2h$ vertices.

Enlarging the window provides a further useful result. Consider $t \cdot h^2 + \binom{x+1}{2}$ consecutive edges in the ordering. The argument in the proof establishes that at most $(t + 1)h + x$ vertices are involved in these edges, and hence:
**Proposition 2.** If $G(h, w)$ has a wrapped $\rho$-labelling, then $c_{t(t+1)h+x}(2wh^2+1) \geq th^2 + \binom{x+1}{2}$ for all $t \geq 1$.

We show some $h \times (w + 1)$ boxes giving wrapped $\rho$-labellings (found by a backtracking algorithm) for $G(h, w)$:

### 3.2 Wrapped $\rho$-Labelling of $\Gamma(h, w)$

A very similar approach enables us to treat windows which induce an odd number of vertices. We form a graph $\Gamma(h, w)$ which contains all of the vertices and edges of $G(h, w)$, along with one further vertex $\infty$, and further edges $\{\infty, v_{ij} \} : 0 \leq i < h, 0 \leq j < w \} \cup \{\infty, v_{iw} \} : 0 \leq i < h \}$. The edge ordering employed for $G(h, w)$ is extended to $\Gamma(h, w)$ by placing all edges involving $\infty$ at the beginning. The property of being wrapped ensures that the vertices $\{\infty\} \cup \{v_{i0} : 0 \leq i < h\}$ have the same labels in $(\Gamma(h, w), \rho)$ as those of $\{\infty\} \cup \{v_{iw} : 0 \leq i < h\}$ in $(\Gamma(h, w), \alpha + \rho)$, so that an identification of these vertices can be made as for wrapped $\rho$-labellings. Figure 4 shows a wrapped $\rho$-labelling of $\Gamma(3, 2)$, adjoined to the first translate with $\alpha = -1$. The vertex labeled 0 is $\infty$ in the first $\Gamma(3, 2)$, while the vertex labeled 48 is $\infty$ in the second.

**Proposition 3.** If $\Gamma(h, w)$ has a wrapped $\rho$-labelling, then there is an ordering of the edges of $K_{2wh(h+1)+1}$ in which every set of $h(h+1)$ consecutive edges induces a subgraph on at most $2h+1$ vertices, i.e. a $(h(h+1), 2h+1)$-cluttered ordering. Thus $m_{h(h+1)}(2wh(h+1)+1) \leq 2h+1$ and $c_{2h+1}(2wh(h+1)+1) \geq h(h+1)$. 

**Fig. 4.** $\Gamma(3, 2)$ and its first translate
This extends in the same manner as Proposition 2 to establish:

**Proposition 4.** If $\Gamma(h, w)$ has a wrapped $\rho$-labelling, then $c_{(t+1)h+x+1}(2wh(h+1)+1) \geq th(h+1) + \left(\frac{x}{2}+1\right)$ for all $t \geq 1$.

In the wrapped $\rho$-labellings to follow, we always take $\rho(\infty) = 0$.

Cohen and Colbourn [7] treated a restricted version of this problem, namely finding wrapped $\rho$-labellings of $\Gamma(1, w)$. They settled this problem completely in the affirmative, and established more. When $d = 3$, the access cost for a specific subgraph is at least 3, and at most 6. When three consecutive edges induce a triangle, we have access cost 3 and this is the only situation accounting for this minimum. However, when $\{e_i, e_{i+1}, e_{i+2}\}$ induces a triangle, $\{e_{i-1}, e_i, e_{i+1}\}$ cannot form a triangle unless $n = 3$, for $e_{i-1} = e_{i+2}$ is necessary if both are triangles. Now when three consecutive edges do not form a triangle, the fewest vertices induced is four, which can be realized by a path or a star on three edges. Hence the minimum 3-access cost of $K_n$ is at least 3.5 when $\binom{n}{2} \equiv 0 \pmod{2}$, and is at least $3.5 - \frac{1}{n^2-n-2}$ when $\binom{n}{2} \equiv 1 \pmod{2}$.

An edge ordering for $K_n$ which realizes this minimum when $d = 3$ is then sought. The specific question addressed is: When can the $\binom{n}{2}$ edges of the complete graph $K_n$ be ordered by a permutation $\pi$, so that among the $\binom{n}{2} - 2$ subgraphs $\{G_i, 3\}$, at least $\frac{1}{4}(n^2 - n - 6)$ subgraphs form triangles? Such an ordering of the edges of $K_n$ is a **ladder ordering of pairs**. When $\binom{n}{2}$ is even, this requires that for any three consecutive edges $e_i, e_{i+1}, e_{i+2}$ with $0 \leq i \leq \binom{n}{2} - 3$, the subgraph induced by these three edges contains three vertices when $i$ is even, and four vertices when $i$ is odd. A wrapped $\rho$-labelling of $G(1, w)$ gives a ladder ordering of pairs for $n = 4w + 1$. Cohen and Colbourn [7] prove:

**Theorem 1.** A ladder ordering of pairs for $K_n$ exists except possibly when $n \in \{15, 18, 22\}$.

This has an immediate application to the determination of maximum access cost:

**Proposition 5.** There is an edge ordering of $K_n$ with $(2f - 4)$-maximum access cost less than or equal to $f$. Hence, with the stated exceptions for $n$, the largest $d$ for which $m_d(n) = f$ is at least $2f - 4$, i.e. $c_f(n) \geq 2f - 4$. 
Proof. A ladder ordering of \( K_n \) employs at most \( f \) vertices among any \( 2f - 4 \) consecutive edges, for all \( f \). It remains only to treat the case when \( f = 3 \), which is handled in \([7]\).

We next develop a simple lower bound for \( c_f(n) \).

**Proposition 6.** For \( f \geq 3 \), \( c_f(n) \leq \binom{f-1}{2} + 1 \). Hence \( c_3(n) = 2 \) and \( c_4(n) = 4 \).

**Proof.** Consider a \((d,f)\)-cluttered ordering \( \pi \) on \( G = K_n \). It must happen that for some value of \( i \), \( G_i^{\pi,d} \) and \( G_{i+1}^{\pi,d} \) are on different sets of \( f \) vertices. But \( G_i^{\pi,d} \) has at most one edge that does not appear in \( G_{i+1}^{\pi,d} \), and hence contains a vertex of degree one. It follows that \( G_i^{\pi,d} \) can contain at most \( \binom{f-1}{2} + 1 \) edges. \( \square \)

This can be refined to prove that \( c_5(n) = 6 \), as follows. If an ordering \( \pi \) of \( K_n \) had 7-maximum access cost equal to five, then by the argument given, we can assume that \( G_i^{\pi,d} \) and \( G_{i+1}^{\pi,d} \) share a \( K_4 \). Upon considering the vertices of \( G_{i+1}^{\pi,d} \) for \( r = 2, 3, 4 \), since no new vertices can be introduced, we find that the three edges \( e_{\pi(i+1)}, e_{\pi(i+2)} \) and \( e_{\pi(i+3)} \) share a common vertex which does not appear in \( G_{i+1}^{\pi,d} \) (otherwise, the vertex \( y \) added in forming \( G_{i+1}^{\pi,d} \) has become saturated and no more edges can be placed).

A more involved argument establishes that \( c_6(n) \leq 9 \), but a general bound is elusive.

### 3.3 A Generalization

There is no wrapped \( \rho \)-labelling of \( G(h,1) \) for \( h \in \{6,7,8,9\} \). However, removing the edge \( \{v_{50}, v_{51}\} \) from \( G_{6,1} \), we can find a wrapped \( \rho \)-labelling of the resulting graph, with \( \rho(v_{00}) = \rho(v_{11}) - 1 = 27, \rho(v_{10}) = \rho(v_{21}) - 1 = 18, \rho(v_{20}) = \rho(v_{31}) - 1 = 41, \rho(v_{30}) = \rho(v_{41}) - 1 = 7, \rho(v_{40}) = \rho(v_{51}) - 1 = 3, \) and \( \rho(v_{50}) = \rho(v_{01}) - 1 = 0 \). By the reasoning of Proposition \([1]\) a \((35,12)\)-cluttered ordering exists for \( K_{7,1} \). To generalize, suppose that we are to form an \((e,2\ell)\)-cluttered ordering of \( K_{2e+1} \) and that \( e \leq \ell^2 \). We could begin with \( G(\ell,1) \), and preserve any \( e \) of the edges. Then if the resulting graph has a wrapped \( \rho \)-labelling, the ordering is easily generated. The only potential difficulty here is in determining which \( e \) edges to retain. Instead, we can label the vertices of \( G(\ell,1) \), not removing any edges, so that every nonzero difference modulo \( 2e + 1 \) is represented as the edge difference on at least one edge; call this a wrapped near-\( \rho \)-labelling of \( G(\ell,1) \) modulo \( 2e + 1 \). If a wrapped near-\( \rho \)-labelling exists, we can omit edges one by one whose difference is multiply represented, until every difference is represented exactly once. Thus we are asking, in general, for wrapped near-\( \rho \)-labellings modulo \( v \) for \( v \leq 2\ell^2 + 1 \). Indeed, there is no requirement that the modulus be odd; when it is even, an edge difference of half the modulus generates two copies of each edge, and hence half the edges of this difference can be omitted and the other half retained. A wrapped near-\( \rho \)-labelling modulo \( 2e + 2 \) yields a cluttered ordering for \( e \) edges. Space does not permit a full exploration here, so we content ourselves with some examples. \( G(7,1) \) has no wrapped \( \rho \)-labelling (modulo 99), but we have found wrapped near-\( \rho \)-labellings for every odd modulus from 51 to 93 inclusive, and for every even modulus from 52 to 88 inclusive. This
establishes that for $25 \leq e \leq 46$, there is an $(e,14)$-cluttered ordering of $K_{2e+1}$, and for $25 \leq e \leq 43$, there is an $(e,14)$-cluttered ordering of $K_{2e+2}$.

The labellings all have a similar structure, with $\rho(v_{i+1,1}) - 1 = \rho(v_{i,0})$ for all $i$, and $\rho(v_{0,0}) = 0$. We therefore present only the labels on $\{x_{i0}\}$ for $0 \leq i \leq 5$ in the following:

<table>
<thead>
<tr>
<th>$v$</th>
<th>$\rho(v_{i0})$</th>
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<th>$\rho(v_{i0})$</th>
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<th>$\rho(v_{i0})$</th>
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<th>$\rho(v_{i0})$</th>
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<tbody>
<tr>
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<td>2 4 6 8 18 29</td>
<td>62</td>
<td>2 4 6 13 30 45</td>
<td>63</td>
<td>2 4 6 8 19 30</td>
</tr>
<tr>
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<td>2 4 6 13 46 31</td>
<td>65</td>
<td>2 4 6 8 19 41</td>
<td>66</td>
<td>2 4 6 14 32 48</td>
<td>67</td>
<td>2 4 6 8 21 32</td>
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<tr>
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<td>2 4 6 11 33 56</td>
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<td>71</td>
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<td>91</td>
<td>2 9 30 68 14 49</td>
<td>93</td>
<td>2 9 14 31 69 50</td>
</tr>
</tbody>
</table>

This strategy extends naturally to $G(h,w)$ and to $\Gamma(h,w)$; when $w > 1$, we insist that when edge $\{v_{i,k}, v_{j,k+1}\}$ is removed, all edges of the form $\{v_{i,\ell}, v_{j,\ell+1}\}$ for $0 \leq \ell < w$ are removed, in order to ensure that the resulting ordering is $(e,2h)$- or $(e,2h+1)$-cluttered.

4 Large Windows

In this section, we examine the opposite extreme, when all but one vertex is induced within a window. An upper bound is easily proved:

**Proposition 7.** $c_{n-1}(n) \leq \lfloor \binom{n}{2} / 2 \rfloor - 1$.

**Proof.** We must establish that when $d = \lfloor \binom{n}{2} / 2 \rfloor$, every ordering $\pi$ has $d$-maximum access cost equal to $n$. Suppose to the contrary that there is an ordering $\pi$ with cost less than $n$. Let $G = K_n$, and consider $G^\pi_0$ and $G^\pi_1$. These graphs must have a different isolated vertex, since at most one edge of $K_n$ fails to appear in either; suppose that 0 and 1 are, respectively, the missing vertices. When $\binom{n}{2}$ is even, all edges appear in one of the two graphs and yet $\{0,1\}$ appears in neither, which is a contradiction. So it remains to examine the case when $\binom{n}{2}$ is odd. Then $\{0,1\}$ must be $e_\pi(\lfloor \binom{n}{2} / 2 \rfloor)$. Now any graph $G^\pi_i$ for $1 \leq i \leq d$ contains the edge $\{0,1\}$ and hence must miss a different vertex. Suppose that $G^\pi_1$ is missing vertex $x$, and $G^\pi_d$ is missing $y$. It cannot happen that $x = 0$, or all occurrences of 0 lie among the last $d$ positions and hence all vertices appear there as well. By the same token, we find that $\{x,y\} \cap \{0,1\} = \emptyset$. Since $G^\pi_0$ misses 0 and $G^\pi_d$ misses $y$, the last edge in the ordering must be $\{0,y\}$; in the same way, the first must be $\{1,x\}$. If $\{x,y\}$ is an edge, there is then no position in which to place it in the ordering. Therefore $x = y$. But then
since $G_{1,n}^{d}$ and $G_{2,n}^{d}$ both miss $x$, the total number of occurrences of $x$ in edges is at most two, the final contradiction.

Now we prove a matching upper bound.

**Proposition 8.** $c_{n-1}(n) \geq \lceil \binom{n}{2}/2 \rceil - 1$.

**Proof.** We define an ordering $\pi$ on the edges of $K_n$. Let 0, 1, 2, and 3 be specific vertices of $K_n$, and let $A$ be the set of remaining $n-4$ vertices. Edges are ordered as follows: $\{2,3\}$, $\{0,2\}$, $\{0,3\}$; $\{3,a\}$ for $a \in A$; $\{0,a\}$ for $a \in A$; $X$; $\{0,1\}$; $Y$; $\{1,a\}$ for $a \in A$; $\{2,a\}$ for $a \in A$; $\{1,2\}$, $\{1,3\}$. Here $X$ and $Y$ partition the remaining edges arbitrarily, subject to the condition that $Y$ contains at least as many edges as does $X$, and contains at most one more edge. It is then easily verified that the graphs $G_{t}^{d}$ with $d = \lceil \binom{n}{2}/2 \rceil - 1$ miss first vertex 1, then 2, then 3, and finally 0.

For $n = 8$, here is an example ordering: $23, 02, 03, 34, 35, 36, 37, 04, 05, 06, 07, 45, 46, 47, 01, 56, 57, 67, 17, 16, 15, 14, 27, 26, 25, 24, 12, 13$. All sets of 13 consecutive edges only involve seven of the vertices. In fact, the only set of 14 consecutive edges involving all eight vertices is the last.

## 5 Inflation

While the techniques of the two preceding sections often give useful cluttered orderings for the RAID application, they can be difficult to apply to specific values of $n$ and $d$. We therefore develop a simple technique, inflation, for providing reasonable orderings in cases where our earlier techniques do not apply directly. Suppose we wish to produce an edge ordering for $K_n$, and we have $(d, f)$-cluttered orderings for a collection of smaller complete graphs. If there is a $(d, f)$-cluttered ordering of $K_t$ for $t < n$, write $n = st + r$, with $0 \leq r < t$. Form a complete graph $K_n$, and partition its vertices into $t$ classes, each having size $s$ or $s+1$. Call the classes $X_1, \ldots, X_s$. For each $1 \leq i \leq s$, the edges within class $X_i$ are partitioned into $t-1$ classes, $I_{i,1}, \ldots, I_{i,t-1}$, so that the number of edges within the edge classes are as equal as possible. Then each has at least $\lceil \binom{|X_i|}{2}/(t-1) \rceil$ edges. Now form a $(d, f)$-cluttered ordering of $K_t$ on vertex set $\{1, \ldots, t\}$. We order the edges of $K_n$ as follows. For each edge $\{i, j\}$ of $K_t$ in turn, we list all edges of $K_n$ with one endpoint in $X_i$ and the other in $X_j$ (in any order). Then if this is the $t$th occurrence of $i$ in an edge, we list the edges in $I_{i,t}$. We repeat this for $j$. The resulting ordering of $K_n$ has `inflated’ every edge of $K_t$ into a collection of at least $s^2 + 2\lceil \binom{s}{2}/(t-1) \rceil$ edges. Hence we have proved:

**Proposition 9.** If a $(d, f)$-cluttered ordering of $K_t$ exists, a $(d', f')$-cluttered ordering of $K_{st+r}$ also exists, where $f' = \min(sf + f, sf + r)$ and $d' \geq (s^2 + 2\lceil \binom{s}{2}/(t-1) \rceil) \cdot d$.

A very special case arises when $n = t + 1$. Then we find that $c_f(n) \geq c_{f-1}(n-1)$. 

Cluttered Orderings for the Complete Graph 429
We do not expect inflation to yield edge orderings which actually minimize the number of vertices induced. Rather, it affords us a method for employing the specific solutions found by combinatorial methods (as in earlier sections) to produce acceptable orderings for a large range of relevant parameters.

6 Concluding Remarks

The impact of improved edge orderings on performance is treated in [8]. In this paper, we have focussed on combinatorial methods for producing cluttered orderings. We extended the notion of ladder orderings (which handle very small reads) to larger reads by establishing the existence of numerous wrapped labellings. The specific examples produced are of value in erasure codes of practical size. We also established an exact result when all but one of the vertices are induced within a window. Finally, we described a simple technique for employing existing orderings for small complete graphs to develop orderings for larger complete graphs; again, this can be expected to have practical value in finding reasonable (but almost certainly not optimal) orderings for particular parameters in a RAID system.

We have treated only the complete graph here, and it is natural to ask about orderings of other graphs as well. In the RAID application, we expect that simply removing the trailing edges from an ordering provides an erasure code with fewer information disks that enjoys similar performance benefits. However, considerations such as reliability and balanced ordering (see [10]) may argue against the omission of those particular edges; hence it may be of interest to extend the techniques here to handle a broader class of graphs.

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References


Improved On-Line Stream Merging: From a Restricted to a General Setting

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Abstract. Stream merging is a promising technique for reducing server bandwidth in video-on-demand systems. There are many heuristics for the problem proposed whose effectiveness has been confirmed empirically. However, it is desirable to prove their effectiveness mathematically. In the pioneering work [2], Bar-Noy and Ladner studied stream merging using competitive analysis. They designed an $O(\log n)$-competitive on-line scheduler, where $n$ is the total number of stream requests. However, their result is applicable only to systems with large client bandwidth and buffer size. In this paper we design the first on-line scheduler for stream merging in the general setting, in which we lift the large resource requirements, and our scheduler achieves constant competitive ratio.

1 Introduction

We study an on-line scheduling problem arising from video-on-demand (VOD) systems that support stream merging. A VOD system often receives many requests for a hot video over a short period of time (say, Friday 7 pm to 9 pm). If the video server uses a dedicated video stream to serve each request, the bandwidth requirement for the server becomes enormous. A simple alternative is to batch [1,5] the requests so that a single stream is issued to serve requests arrived over a period of time; however, this increases the response time substantially.

To reduce the bandwidth requirement without sacrificing the response time, Eager et al. [6,7] proposed a promising technique called stream merging, which assumes that a client is equipped with extra buffer and can receive data from two streams simultaneously. The technique takes advantage of the assumptions as follows. Initially, a client receives data from a dedicated stream $X$ for immediate playback. After some time, the client may start to buffer data simultaneously from another stream $Y$ initiated earlier (due to another client). Suppose $Y$ is initiated $t$ time units before $X$. Then after listening to both $X$ and $Y$ for $t$ time units, the client no longer needs $X$ as it can play back the next $t$ time units using the data in the buffer, while further buffering data from $Y$ simultaneously. Hence, $X$ may terminate now. We say that $X$ merges with $Y$ when the client starts to count on $Y$ only. Merging streams allows several clients to be served using only one full stream plus a number of partial streams, thus saving the bandwidth of the server. Figure 1 shows an example.

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Fig. 1. (a) At time 0, client B is served by stream Y. (b) At time 1, client A arrives. It is served by stream X; it simultaneously buffers data from Y. (c) At time 2, A plays the content of its buffer and continues buffering from Y. X terminates. Assume the video length is 120 time units. The required bandwidth is 121 instead of 240.

An on-line scheduler for stream merging determines whether each existing stream should continue to run or merge with another earlier stream without the knowledge about the requests not yet arrived. The goal is to minimize the total bandwidth used, i.e., the total duration of all streams. Such schedulers were first studied in a primitive form in [3,9,10], in which a stream is allowed to merge with a full stream only. Eager et al. [6–8] gave a general on-line scheduler that performs well empirically. Bar-Noy and Ladner [2] was the first to study the problem using competitive analysis. An on-line scheduler is said to be $c$-competitive, or has competitive ratio at most $c$, if for any sequence $I$ of requests, it uses at most $c$ times of the bandwidth used by the optimal off-line schedule in serving $I$. In [2], Bar-Noy and Laner gave an on-line scheduler with competitive ratio at most $\min\{\log_\phi n, \log_\phi \frac{1}{2D}\} + O(1)$, where $\phi = (1 + \sqrt{5})/2$, and $D < 1$ is the guaranteed startup delay measured as a fraction of the time for a full stream. In [4], we gave the first on-line scheduler with constant competitive ratio.

All the above-mentioned results assume that the total client bandwidth is at least twice the playback bandwidth (half of it is for normal playback and the additional half is for buffering an earlier stream) and the client buffer is sufficient to store up to half the size of a video. Eager et al. [8] pointed out that in reality, clients may have limited bandwidth and buffer. For the general problem—the $(\omega, B)$-stream merging problem, in which the client additional bandwidth is only a fraction $\omega \ (0 < \omega \leq 1)$ of the playback bandwidth and the client can buffer at most $B \ (B > 0)$ time units of video, only empirical results are presented (see [8]). In [2], Bar-Noy and Ladner left as an open problem whether the requests can be handled competitively in this general setting.

In this paper we give the first on-line scheduler for $(\omega, B)$-stream merging with constant competitive ratio (precisely, 5) for any given $0 < \omega \leq 1$ and any $B > 0$, and thus solving the open problem of Bar-Noy and Ladner. Note that the competitiveness of our scheduler does not deteriorate even if the bandwidth and buffer are limited. Our result is based on a novel observation that the behavior
of any scheduler can be modeled by rectilinear (binary) trees on a triangular grid, and we will give the details in Section 2. In Section 3, we present a general methodology for constructing a competitive on-line scheduler from some near-optimal rectilinear tree. In Section 4, we show how to construct such near-optimal rectilinear tree.

2 Concepts and Notations

2.1 Preliminaries

In our discussion, time is divided into intervals of unit length, which can be arbitrarily small. Requests arriving within the same interval are treated as a single request and they are served by multicasting one stream. We suppose both video length $L$ and buffer size $B$ are measured in terms of time units.

Note that with the additional bandwidth being only a fraction $\omega$ of the playback bandwidth, it takes longer time for a stream $X$ to merge with a stream $Y$ initiated $t$ time units earlier. It is not difficult to see that $X$ takes $\lceil 1/\omega \rceil t$ time units to buffer data from $Y$ before the two streams can merge (see [8]). We call $C = \lceil 1/\omega \rceil$ the catch up ratio. Observe that for $X$ merging to $Y$ successfully, the buffer size $B$ could be equal to $t$, and since $X$ must merge with $Y$ before $Y$ terminates, we require $t + Ct \leq L$, or equivalently $t \leq L/(1+C)$. Thus, $B$ can be as small as $t \leq L/(1+C)$. On the other hand, if $B$ is smaller than $L/(1+C)$, the optimal off-line schedule has only limited choices to merge the streams, and an on-line scheduler can simulate the off-line schedule easily (details given in the full paper). Thus, we only need to consider the case where $B = L/(1+C)$.

A request sequence is said to be compact if the last request arrives within $B$ time units from the first request. Notice that such a sequence contains at most $B + 1$ requests. Intuitively, all streams initiated for a compact request sequence can merge successfully before the first stream terminates, and the first stream initiated must be a full stream. In fact, using an observation of Bar-Noy and Ladner [2], we can show that for compact request sequences, it suffices to focus on schedules with the following property:

The first stream is the only full stream and all other streams merge with other streams eventually.

Since the first stream always contributes a dominating portion in the overall bandwidth usage of any schedule, the overall bandwidth of different schedules may not differ a lot. To make a more meaningful comparison between different schedulers, we define the full cost to be the bandwidth used by all the streams, and define the merge cost to be the bandwidth used by all except the first stream. Note that the full cost and merge cost differ by exactly the video length $L$.

Compact request sequence is a key notion for stream merging because a general request sequence can be processed competitively if each consecutive compact request sequence can be processed competitively. More precisely, suppose $A$ is an on-line scheduler for the problem such that with respect to compact request
sequences, its merge cost is at most \(c\) times that of the optimal off-line schedule. We can use \(\mathcal{A}\) to process arbitrary request sequence as follows. As requests arrive, we divide them into groups. When a request arrives within \(B\) time units from the first request of the current group, it is inserted to this group, and we use \(\mathcal{A}\) to schedule a stream for it and determine how the stream merges with other streams. Otherwise, we start a new group and initiate a full stream for this request. Notice that each group of requests is in fact a compact request sequence. The following theorem, which generalizes the one in [2], asserts that the bandwidth used by the resulting schedule is at most \(c + \frac{\mathcal{L}}{(\mathcal{L} - B)}\) times that of the optimal off-line schedule. The proof will be given in the full paper.

**Theorem 1.** Let \(\mathcal{A}\) be an on-line scheduler such that with respect to compact request sequence, its merge cost is at most \(c\) times that of the optimal off-line schedule. Then for any arbitrarily long request sequence, using \(\mathcal{A}\) to schedule separately each compact request sequence can produce a schedule with an overall bandwidth at most \(c + \frac{\mathcal{L}}{(\mathcal{L} - B)}\) times that of the optimal off-line schedule.

In the rest of this paper, we focus on compact request sequences and give an on-line scheduler whose merge cost is at most three times that of the optimal off-line schedule. Then, by Theorem 1 the competitive ratio of this scheduler is at most \(3 + \frac{\mathcal{L}}{(\mathcal{L} - B)} = 5\) because \(\mathcal{C} \geq 1\), and we are consider the case \(B = \mathcal{L}/(1 + \mathcal{C})\). Unless specified otherwise, all request sequences mentioned are meant to be compact request sequences. To ease discussion, we assume that the first request of all compact request sequences arrives at time 0.

### 2.2 Binary Merge Trees

In this section we show how to model schedules using binary merge trees, which attempt to model the occurrences of merging and the status of each stream.

Below, we use the notation \(r(s)\) to denote the time when a stream \(s\) is initiated.

A binary merge tree is a rooted, ordered binary tree. For any node with (out-)degree one, we require the node to have a left son but no right son. Every node \(u\) is associated with a stream \(s_u\) and a time \(t_u\). In particular, every leaf \(\ell\) corresponds to a distinct stream that is initiated due to a request, and \(t_\ell = r(s_\ell)\).

For an internal node \(u\), its associated time is always greater than that of its sons, and its stream is always inherited from its left son. If \(u\) is of degree two, \(u\) and its right son, denoted by \(v\), represent the occurrence of a merging. We require that \(r(s_v) > r(s_u)\) and \(t_u = t_v + \mathcal{C}(r(s_v) - r(s_u))\). At time \(t_v\), the stream \(s_v\) (precisely, the clients served by \(s_v\)) starts to buffer the content of \(s_u\); after \(\mathcal{C}(r(s_v) - r(s_u))\) time units (i.e., at time \(t_u\), \(s_v\) merges with \(s_u\) and terminates.

Regardless of whether \(u\) is degree one or two, \(u\) and its left son \(w\) corresponds to a continuation of the same stream from time \(t_w\) to time \(t_u\). That is, the stream has no plan to merge with another earlier stream during this interval. Note that the root as well as all the nodes on the leftmost path of the tree are associated

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1 This is in contrast to the merge trees used by Bar-Noy and Ladner [2], which model the relationship among the streams.
with the same stream, which does not merge with any stream and runs till the
end. The time associated with the root is exactly $L$.

The cost of an edge $(u, v)$ is defined as $t_u - t_v$. The full cost of a binary
merge tree is the sum of the cost over all edges, and the merge cost is the full
cost minus the cost of the edges on the leftmost path.

Recall that we focus on compact request sequences and schedules in which all
streams merge with the first one. Note that each binary merge tree corresponds
to one of these schedules with the same full cost and merge cost, and vice versa.
Thus, for any compact request sequence, there is a binary merge tree with full
cost and merge cost equal those of the optimal off-line schedule. Below, we show
that the optimal binary merge tree, the tree associated with an optimal off-line
schedule, can be assumed to have a more uniform structure.

**Lemma 1.** Consider any compact request sequence $I$. There is an optimal bi-
nary merge tree for $I$ such that all the internal nodes, except possibly the root,
have both left and right sons.

**Proof.** Let $T$ be the optimal binary merge tree with the minimum number of
nodes. Assume there is a degree one internal node $v$. By definition, $v$ has a left
son, say $w$. Let $u$ be the parent of $v$. If $v$ is the left son of $u$, we remove $v$ and add
an edge between $u$ and $w$. The cost is not changed; contradicting that $T$ has
the minimum number of nodes. If $v$ is the right son of $u$, $s_v$ starts buffering from $s_u$
at $t_v$. We make $s_v$ starts buffering from $s_u$ earlier, say at $t_w$. This decreases the
merge cost of the schedule and contradicts the optimality of $T$. $\square$

### 2.3 Rectilinear Trees

This section discusses a powerful tool called *rectilinear trees* to model schedules
of stream merging. The new tool is more useful for visualizing the difference
between an arbitrary schedule and the optimal off-line schedule. Roughly speak-
ing, rectilinear trees are binary merge trees embedded on a triangular grid. Let
$K = B + 1$. Consider a $K \times K$ grid with $K$ grid points (i.e., with $K - 1$ grid
arcs) on each boundary. The grid is distorted into a rectangle: The horizontal
grid arc has a length of $(1 + C)$ and the vertical grid arc a length of $C$.

A grid line is a sequence of consecutive horizontal (or vertical) grid arcs. A grid path is a sequence of grid arcs connecting any two grid points. The rows
and columns of the grid are numbered 0 to $K - 1$ from top to bottom, and
from left to right, respectively. The triangular grid in concern is the upper right
triangle with grid points $(0,0), (1,1), \cdots, (K-1, K-1)$ as the hypotenuse.

A rectilinear tree is a binary tree on the triangular grid satisfying the follow-
ing properties. (1) The root is at the top rightmost corner of the triangular grid
and the leaves are on the hypotenuse of the triangular grid. (2) Any node and
its parent are either on the same row or on the same column, i.e., any edge is
either a horizontal or a vertical grid line. (3) The grid path from any leaf to the
root is monotone—The rectilinear distance to the root is strictly decreasing as
we traverse from the leaf to the root.
For any rectilinear tree $T$, we define the length of an edge to be the length of the corresponding horizontal or vertical grid line. The full length of $T$ is the total length of all its edges, and the branch length is the total length of all edges except those edges along the leftmost path of the tree. Notice that two rectilinear trees with the same set of grid arcs have the same branch length even if the set of nodes may differ. Branch length is a notion heavily used below; to simplify our discussion, we denote by $\delta(T)$ the branch length of $T$. The following theorem shows the correspondence between binary merge trees and rectilinear trees.

**Theorem 2.** Any binary merge tree can be embedded on a triangular grid as a rectilinear tree with branch length equal to the merge cost of the binary merge tree.

**Proof.** Let $T$ be a binary merge tree. For any node $u$ in $T$, we create a node $u_R$ at the grid point $(r(s_u), (t_u + Cr(s_u))/(1+C))$. If $u$ is a leaf, $u_R$ is on the hypotenuse. The root becomes a node at the upper rightmost corner of the grid. If $u$ is an internal node, let $w$ be the left son of $u$. $w_R$ is at $(r(s_w), (t_w + Cr(s_w))/(1+C))$. As $s_u$ is the same as $s_w$ and $t_u > t_w$, $u_R$ lies on the right of $w_R$ on the same row. Thus, the edge $(u, w)$ becomes a horizontal grid line between $u_R$ and $w_R$, with length $(1+C) \times ((t_u + Cr(s_u) - t_w - Cr(s_w))/(1+C) = t_u - t_w$, which equals the cost of the edge $(u, w)$ in $T$. The case for the right son can be proved similarly. The branch length of the rectilinear tree equals the merge cost of $T$. $\square$

### 3 The On-line Scheduler

A crucial issue in designing on-line scheduler is to estimate the cost of the optimal solution. We give a simple way to estimate the cost of the optimal solutions of all compact request sequences using only one rectilinear tree, called the complete rectilinear tree. This rectilinear tree when restricted to any compact request sequence $I$ gives an attainable estimate of the branch length of the optimal rectilinear tree for $I$. Our on-line scheduler makes reference to this tree to schedule the merging of streams; the resulting schedule corresponds to a rectilinear tree with branch length always upper bounded by the estimated branch length of the optimal rectilinear tree.

Define a complete rectilinear tree $T$ to be a rectilinear tree whose leaves comprise all the grid points along the hypotenuse of the triangular grid. For any compact request sequence $I$, for simplicity, $I$ also denotes the subset of grid points along the hypotenuse corresponding to the requests. The rectilinear tree $T$ when restricted to $I$ is called the partial rectilinear tree of $T$ with respect to $I$. It is a rectilinear tree denoted by $T_{\parallel I}$ whose set of leaves is $I$ and set of grid arcs is a subset of that of $T$. In other words, for any grid arc $e$ in $T$, if $e$ is along a path of $T$ from the root to some leaves in $I$, $e$ is in $T_{\parallel I}$; otherwise, $e$ is not in $T_{\parallel I}$. Note that $T_{\parallel I}$ is unique for any given $I$.

In this section, we describe a methodology which makes reference of a given complete rectilinear tree $T$ to derive an on-line algorithm $A_T$. For any compact request sequence $I$, $A_T$ gives a schedule whose corresponding rectilinear tree, denoted by $A_T(I)$, has a branch length at most that of $T_{\parallel I}$. That means for
any given $T$ and $I$ we can produce a schedule with merge cost at most $\delta(T \parallel I)$. On the other hand, there is an optimal binary merge tree for $I$ satisfying the property stated in Lemma 1. By Theorem 2, this tree can be embedded on the triangular grid as a rectilinear tree, denoted by $OPT_I$. In Section 4 we describe a particular complete rectilinear tree $S$ such that $\delta(S \parallel I) \leq 3\delta(OPT_I)$, and hence the merge cost of the schedule produced by $A_S$ is at most three times that of the optimal off-line schedule. Together with Theorem 1 we have the next theorem.

**Theorem 3.** $A_S$ is 5-competitive.

In this section we assume $T$ is a chosen complete rectilinear tree. We give an on-line scheduler $A_T$ which constructs a rectilinear tree based on $T$ and a compact request sequence. The scheduler makes reference to $T$ to determine how the streams serving the requests merge with each other. Details of $A_T$ are as follows.

When the first request arrives (at time 0), the scheduler initiates a stream for this request and this stream will not merge with any stream. The initial rectilinear tree contains only two nodes, namely, the root at the top rightmost corner and a leaf at the top leftmost corner, of the triangular grid. These two nodes are connected by a horizontal grid line. All the grid points along this grid line are associated with the first stream.

When a subsequent request arrives, it is assigned a dedicated stream $X$ immediately. We also determine when $X$ should begin buffering data from which earlier stream. Let $(r, r)$ denote the leaf corresponding to current request where $r$ is its arrival time. We locate in $T$ the least common ancestor $a$ of $(r, r)$ and the leaf corresponding to the request just before the current request. A new grid path is added to the current rectilinear tree as follows.

If $a$ lies on column $r$, the new grid path is a vertical grid line running upwards from $(r, r)$. The path stops at the first grid point it intersects with the current rectilinear tree. In this case, $X$ immediately buffers data from the stream associated with the intersecting grid point.

If $a$ does not lie on column $r$, it must lie on the right of column $r$ because of the monotone property of the paths in $T$. Suppose $a$ lies on column $c_a$. The new grid path comprises a horizontal grid line from $(r, r)$ to $(r, c_a)$ and a vertical grid line from $(r, c_a)$ to the first grid point it intersects with the current rectilinear tree on column $c_a$. In this case, $(1 + C)(c_a - r)$ time units after $X$ is initiated, $X$ starts buffering from the stream associated with the intersecting grid point.

All the grid points along the new path, except the intersecting point, are associated with $X$. The decision is made every time a request arrives. For any compact request sequence $I$, the next lemma bounds $\delta(A_T(I))$ by $\delta(T \parallel I)$.

**Lemma 2.** For any complete rectilinear tree $T$ and compact request sequence $I$, $\delta(A_T(I)) \leq \delta(T \parallel I)$.

**Proof.** (Sketch.) Suppose $I$ has $k$ requests. For $1 \leq i \leq k$, let $I_i$ be the first $i$ requests in $I$, and $A_T^{i}$ be the rectilinear tree constructed by $A_T$ for $I_i$. We claim that $A_T^{i}$ is always on the right of $T \parallel I_i$, precisely, for any grid point $(r, c) \in T \parallel I_i$,
Fig. 2. Suppose \( C = 1 \). A \( 12 \times 12 \) triangular grid \( G \) is divided into an \( 8 \times 8 \) and a \( 4 \times 4 \) triangular grids \( G_1 \) and \( G_2 \). (a) The complete rectilinear tree \( S \) on \( G \). The unshaded area is the rectangle \( R \) with top rightmost and the bottom leftmost corners at \((0,11)\) and \((8,7)\). (b) An optimal rectilinear tree for a given set of compact request sequence with seven requests. (c) After a recursive step the grid arcs within the rectangle (unshaded area) are removed and the grid arcs on the four boundaries are included.

there exists a grid point \((r',c)\) \( \in A^1_T \) with \( r' \geq r \). Since the path in \( T \) is monotone, the grid path added to \( A^1_{T-1} \) for the \( i \)-th request has a length at most that of the path added to \( T \parallel I_{i-1} \) to form \( T \parallel I_i \). Thus, \( \delta(A_T(I)) \leq \delta(T \parallel I) \) as they are the sums of the corresponding path lengths. \( \square \)

4 A Good Complete Rectilinear Tree

In this section we describe a complete rectilinear tree \( S \) on a \( K \times K \) triangular grid \( \Delta \), which provides a good estimate of all optimal rectilinear trees. Precisely, we show that \( \delta(S \parallel I) \leq 3 \delta(OPT_I) \) for all compact request sequences \( I \).

We construct \( S \) recursively. For any integer \( n \), let \( f(n) = \lfloor \frac{(1 + C)}{1 + 2C} \right(n - 1) + 1 \). The root of \( S \) is at \((0,K-1)\) of \( \Delta \). All the \( K \) grid points along the hypotenuse are the leaves of \( S \). If \( S \) has only one leaf (i.e., when \( K = 1 \)), we are done. If \( S \) has at least two leaves, it has two subtrees. The left subtree has its root at \((0,f(K)-1)\) and it contains the top \( f(K) \) leaves along the hypotenuse. The right subtree has its root at \((f(K),K-1)\) and it contains the rest \( K - f(K) \) leaves. These two subtrees are then constructed recursively, and their roots are connected to the root of \( S \) by a horizontal grid line with \( K - f(K) \) grid arcs and a vertical grid line with \( f(K) \) grid arcs. See Figure 2(a).

To prove \( \delta(S(I)) \leq 3 \delta(OPT_I) \), we derive a procedure TRANSFORM that transforms \( OPT_I \) into a rectilinear tree \( O_I \) that uses only grid arcs of \( S \), and whose set of leaves containing \( I \). TRANSFORM is conservative in replacing the edges of \( OPT_I \) so that \( \delta(O_I) \leq 3 \delta(OPT_I) \). On the other hand, \( S \parallel I \) contains the minimal subset of grid arcs in \( S \) to form a rectilinear tree with \( I \) as leaves and using grid arcs of \( S \) only. Thus, \( \delta(S \parallel I) \leq \delta(O_I) \), and \( \delta(S \parallel I) \leq \delta(O_I) \leq 3 \delta(OPT_I) \).

4.1 The Transformation

TRANSFORM is recursive; it takes two parameters \( G \) and \( T \). \( G \) is a triangular grid, and \( T \) is a rectilinear tree embedded on \( G \) whose root is at the top rightmost corner of \( G \). Initially, \( G = \Delta \) and \( T = OPT_I \). In the first step of TRANSFORM, consider the rectangle \( R \) whose top rightmost corner is at \((0,K-1)\) and bottom leftmost corner is at \((f(K),f(K)-1)\) \( \square \) See Figure 2(a). Note that \( S \) does not

\(^2\) \( R \), excluding its bottom leftmost corner, is enclosed entirely in \( \Delta \).
cross $R$; more precisely, no path in $S$ enters $R$ from a boundary of $R$, goes through at least one of its interior grid points, and leaves from a distinct boundary. Thus, we must remove all grid arcs of $OPT_I$ within $R$.

If $OPT_I$ has a path crossing $R$ from the bottom boundary, we expand $OPT_I$ to include all the arcs on the bottom and right boundaries of $R$, except the leftmost one (which is outside $\Delta$). If $OPT_I$ has a path crossing $R$ from the left boundary, we include arcs on the left and top boundaries of $R$, except the bottommost one. We then remove all grid arcs that are within $R$. See Figure 2(b) and (c). New nodes are created on any new intersections of grid arcs introduced.

Let $T_0$ be the new tree formed after the above step. Ignoring the interior grid points of $R$, we partition $\Delta$ into two disjoint triangular grids: an $f(K) \times f(K)$ one $G_1$ to the left of $R$ and a $(K - f(K)) \times (K - f(K))$ one $G_2$ below $R$. See Figure 2(c). Consider the part of $T_0$ in $G_1$. All paths in $T_0$ starting from the leaves in $G_1$ to the root of $T_0$ exit $G_1$ through its top rightmost corner. If this corner is not a node of $T_0$, we add a node there. Then, the part of $T_0$ in $G_1$ forms a tree $T_1$ that is embedded on $G_1$ with the root at its top rightmost corner. If $G_1$ is at least $3 \times 3$ and $T_1$ contains at least a leaf, we recursively call Transform on $G_1$ and $T_1$ to form a tree $T_1'$ that uses only grid arcs of $S$ and retains the corresponding leaves. We handle the part of $T_0$ in $G_2$ similarly, and let $T_2'$ be the resulting tree. Note that $T_1'$ and $T_2'$, together with the tree edges of $T_0$ on the boundaries of $R$, form $O_t$.

### 4.2 Analysis of Procedure Transform

In this section we prove that $\delta(O_t) \leq 3\delta(OPT_I)$. Let $E_a$ and $E_r$ be the set of grid arcs added and removed by calling Transform on $G = \Delta$ and $T = OPT_I$. For any set of grid arcs $E$, we denote by $|E|$ the total length of grid arcs in $E$. We first show that $|E_a| \leq 3|E_r|$. Note that it suffices to show that in each recursive step (excluding the recursive calls), the total length of grid arcs added is at most three times that removed. Since the analysis is the same for every recursive step, we only consider the first step, in which the grid arcs in $OPT_I$ within $R$ are removed, where $R$ is the rectangle whose top rightmost and bottom leftmost corners are at $(0, K - 1)$ and $(f(K), f(K) - 1)$ respectively. Let $E_a$ and $E_r$ denote the set of grid arcs removed and added in this particular step respectively.

**Lemma 3.** If $OPT_I$ crosses $R$, $OPT_I$ contains a path crossing $R$ vertically or horizontally, i.e., the path passes through a single column or a single row of $R$.

**Proof.** Suppose $OPT_I$ has a path $\sigma$ from a leaf to the root that crosses $R$. $\sigma$ exits $R$ through either the right or top boundary. If $\sigma$ exits through the right boundary, say on row $i$, $\sigma$ or another path in $OPT_I$ must go through $R$ horizontally on row $i$. Otherwise, there is a node in $OPT_I$ having at most one son. This contradicts Lemma 1 and Theorem 2. The other case can be proved similarly. □

If $OPT_I$ does not cross $R$, obviously $|E_a| \leq 3|E_r|$; otherwise, we have two cases depending on whether $OPT_I$ has a path crossing $R$ vertically or horizontally.

Let $r(K)$ and $t(K)$ denote the length of the right and top boundaries of $R$ respectively. In other words, $r(K) = \mathcal{C}f(K)$ and $t(K) = (1 + \mathcal{C})(K - f(K))$. 

Lemma 4. If $\text{OPT}_1$ contains a path crossing $R$ vertically, $|\mathcal{E}_r| \geq r(K)$ and $|\mathcal{E}_a| - |\mathcal{E}_r| \leq r(K) + t(K) - 1 - 2C$. If $\text{OPT}_1$ contains a path crossing $R$ horizontally, $|\mathcal{E}_r| \geq t(K)$ and $|\mathcal{E}_a| - |\mathcal{E}_r| \leq r(K) + t(K) - 1 - 2C$.

Proof. We only prove the first scenario, and the second one can be proved symmetrically. Let $H_{\text{top}}$ and $H_{\text{in}}$ be the sets of horizontal grid arcs in $\text{OPT}_1$ that are on the top boundary of $R$ and inside $R$, respectively. Thus, $|\mathcal{E}_r| \geq r(K) + |H_{\text{in}}|$.

Suppose $\text{OPT}_1$ has a path crossing $R$ from the left boundary at grid point $p$. TRANSFORM expands $\text{OPT}_1$ to include all four boundaries, excluding the two arcs adjacent to the bottom leftmost corner. Thus, $|\mathcal{E}_a| \leq 2r(K) + t(K) + (t(K) - |H_{\text{top}}|) - 1 - 2C$. Consider the path $\sigma$ from $p$ to the root of $\text{OPT}_1$. The total length of horizontal grid arcs in $\sigma$ is at least $t(K)$, and is at most $|H_{\text{top}}| + |H_{\text{in}}|$. This implies $t(K) \leq |H_{\text{top}}| + |H_{\text{in}}|$. Therefore, $|\mathcal{E}_a| - |\mathcal{E}_r| \leq 2r(K) + t(K) + (t(K) - |H_{\text{top}}|) - 1 - 2C - (r(K) + |H_{\text{in}}|) \leq r(K) + t(K) - 1 - 2C$. The case for $\text{OPT}_1$ not crossing $R$ from the left boundary can be proved easily. □

By simple arithmetic, we have $r(K) \geq t(K) - 1 - C$ and $t(K) \geq r(K) - C$. By Lemma 4, $|\mathcal{E}_a| \leq 3 |\mathcal{E}_r|$. Thus, $|E_a| \leq 3 |E_r|$. Since $E_a$ and $E_r$ are disjoint, $|O_1| = |\text{OPT}_1| - |E_a| - |E_r|$. Note that $\delta(\text{OPT}_1) = |\text{OPT}_1| - \ell$ and $\delta(O_1) = |O_1| - \ell$, where $\ell$ is the length of the top boundary of $\Delta$. Therefore, $\delta(O_1) = \delta(\text{OPT}_1) + |E_a| - |E_r|$. Since $|E_r| \leq \delta(\text{OPT}_1)$, $\delta(O_1) \leq \delta(\text{OPT}_1) + 2|E_r| \leq 3\delta(\text{OPT}_1)$.

Lemma 5. $\delta(O_1) \leq 3\delta(\text{OPT}_1)$.

Recall that $\delta(S||I) \leq \delta(O_1)$. By Lemmas 2 and 5, we have $\delta(A_S(I)) \leq \delta(S||I) \leq \delta(O_1) \leq 3\delta(\text{OPT}_1)$ and we have the following theorem.

Theorem 4. Given any compact request sequence $I$, $\delta(A_S(I)) \leq 3\delta(\text{OPT}_1)$.

References


On-Line Deadline Scheduling on Multiple Resources

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Abstract. We consider an on-line non-preemptive deadline scheduling on multiple resources. Each job has the slack time during which it can wait until its expiration time, the last time it can be started to be completed by the deadline. There are a number of resources on which the jobs are allocated. The effect of slack times on the on-line deadline scheduling was investigated in [8] by analyzing the performance of the Earliest Expiration Time First (EXF) algorithm. But most of the previous work is restricted to a single resource, and in this paper, we propose a different method of analysis which can also be applied to multiple resources. So we can extend the result of [8] to multiple resources.

1 Introduction

Consider the following problem: we are given $m$ resources and a set of job requests, where each job can be served by at most one of the resources and has the processing time for which it uses the resource and the deadline by which it should be completed, and each resource serves only one job at any time. Not all jobs are required to be scheduled and some jobs can be rejected but they cannot contribute to the throughput. Whenever a resource is idle, scheduling algorithms should determine which job to be processed on the resource so that it maximizes the total of gains of accepted jobs.

The schedule is non-preemptive, that is, once a job is scheduled, it cannot be preempted or rejected and it is on-line, where the algorithm has no knowledge of jobs until they arrive and when they arrive, the algorithm is informed the processing times and deadlines of the jobs. The performance of an on-line algorithm is estimated through competitive analysis [9, 12], in which it is compared with that of an off-line optimal algorithm that knows all informations of the future requests.

In this paper, the problem is described in terms of machine scheduling. Each job $J_i$ has the release time $r_i$, the processing time $p_i$, and the deadline $d_i$ and there are $m$ identical machines on which jobs are scheduled. The expiration time $x_i$ of a job $J_i$ is defined to be the last time at which $J_i$ can be scheduled to completion by the deadline, that is, $d_i = x_i + p_i$ and the slack time $s_i$ of $J_i$ to be $x_i - r_i$. If the slack time of a job is more than zero, before scheduling it, an on-line scheduling algorithm has a possibility of waiting for better jobs.
which will arrive. Assume that for each job $J_i$, its slack time satisfies $s_i \geq \kappa p_i$ for some constant $\kappa$. Depending on the value of $\kappa$, we will investigate the effect of the slack time on the performance of the on-line scheduling algorithm. In [8], Goldwasser proposed the Earliest EXpiration Time First (EXF) algorithm that schedules the job with the smallest expiration time among all available jobs at all times. His work is restricted to a single machine, as in the previous results [7] and [10]. In [7] and [8], the authors divide the time busy periods during which there is no idle time on the machine in the schedule of EXF. Since jobs scheduled within a busy period have no effect on the schedules of the other busy periods, it is sufficient to analyze the performance of EXF only in a busy period. But this cannot be applied to multiple machines, where busy periods on distinct machines are aligned to cross at some times. We provide a different analysis that can also show the performance of EXF on multiple machines, where the time is divided into periods in a different way according to the behavior of EXF.

Let $S$ be a set of job instances. Then $EXF(S) (OPT(S))$ represents the sum of the processing times of jobs scheduled by EXF (OPT, resp.) in $S$. Interchangeably, for a time interval $T$, $EXF(T) (OPT(T))$ also represents the set of all jobs scheduled by EXF (OPT, resp.) in $T$. Also $|S|$ is the number of jobs in $S$.

1.1 Previous Work

Off-line non-preemptive scheduling of jobs with deadlines is well-known in the literature from the early days [6, 11], but very recently, there have been active studies for this problem [1, 2], where the jobs have arbitrary release times and weights and the types of machines are given as the single machine, identical machines, and unrelated machines. They give constant approximation algorithms using various optimization techniques, for example, LP-rounding, local search, and etc.

In the on-line setting, it originates from the work of Lipton and Tomkins [10], where each job must be scheduled or rejected as soon as it arrives. For the general case in which the jobs have arbitrary processing times, they give $O((\log \Delta)^{1+\epsilon})$-competitive algorithm and show that no $O(\log \Delta)$-competitive algorithm can exist, where $\Delta$ is the ratio of largest to smallest processing time. In [7], they allowed the jobs to specify arbitrary slacks, i.e., $\kappa \geq 0$, and the effect of the positive slack times was shown in [8]. For equal processing time jobs, he proves EXF is $(1 + \frac{1}{\lfloor \kappa \rfloor + 1})$-competitive and shows that no on-line algorithm can have a better ratio, and for arbitrary processing time jobs, he gives a $(2 + \frac{1}{\kappa})$-competitive algorithm and shows that this is best possible for all deterministic on-line algorithms. But these results are all restricted to the model of a single machine.

For on-line preemptive scheduling problem, there are several related studies [3, 4, 5]. In [4], the effect of slack times for preemptive scheduling on a single machine is investigated, and in [5], the performance of preemptive scheduling on multiple machines is characterized by the stretch factor which is the ratio of the response time to the processing time, that is, $\frac{s_i + p_i}{p_i}$.
2 Equal Processing Time Jobs

2.1 Case \([\kappa] \geq 1\)

Let \(J\) be the set of all job instances. W.l.o.g., we assume each job has the processing time 1. First we consider the case \([\kappa] \geq 1\). According to the behavior of EXF, we divide the time into periods and classify the jobs in \(J\). Here, observe the behavior of EXF. When a new job arrives, EXF puts it into a queue, and whenever there is an idle time on some machine, EXF selects the job with the earliest expiration time in the queue and schedules it on the idle machine. If a job has not been selected until its expiration time, it is rejected at that time.

Inductively, we define the time periods \(T_i = (t_{bi}, t_{bi} - 1)\), \(i = 0, 1, \cdots\), and \(t_{b_{-1}} = \infty\). Assume that the time periods \(T_j, j = 0, \cdots, i\) have been defined and there is a job which EXF rejects in \((s_i, t_i)\). Then we consider the job which is rejected in the last in \((s_i, t_i)\) and denote it by \(I_i\). Let \(t_i\) be the time at which \(I_i\) is rejected and let \(s_i\) be the release time of \(I_i\). Then any job with expiration time > \(t_i\) cannot be scheduled in \([s_i, t_i]\), because if such a job were scheduled, it would have an expiration time greater than \(I_i\) and EXF would schedule \(I_i\). Thus if a job has an expiration time > \(t_i\), either it is scheduled after \(t_i\) or before \(s_i\). We refer to the time \(t_i\) as the critical time, and we will also define the time \(t_{bi+1}\), called the barrier time. Let \(\tilde{t}\) be the time at which an idle time last appears on some machine before \(s_i\). If there is no job with expiration time > \(t_i\) which is scheduled in \([\tilde{t}, s_i]\), we set the time \(t_{bi+1}\) to be \(\tilde{t}\). Otherwise, let \(\hat{I}\) be the job with expiration time > \(t_i\) which is last scheduled in \([\tilde{t}, s_i]\) and let \(\hat{t}\) be the time when \(\hat{I}\) is scheduled. Then at the time \(\hat{t}\), there is no job with expiration time ≤ \(t_i\) in the queue, because if such a job were in the queue, EXF would schedule it in place of \(\hat{I}\). Consider the job, denoted by \(\hat{I}\), which is first scheduled after \(\hat{t}\), and let \(\hat{t}\) be the time at which \(\hat{I}\) is scheduled. Then we can see that new jobs have been released in \((\hat{t}, \hat{t}]\), because otherwise, if \(\hat{t} \geq s_i\), then it is a contradiction to the fact that \(I_i\) was released at \(s_i\), and if \(\hat{t} < s_i\), then \(\hat{I}\) were also in the queue at \(\hat{t}\), that is, \(\hat{I}\) had the expiration time > \(t_i\), and it is a contradiction to the definition of \(\hat{I}\). Thus we set \(t_{bi+1}\) to be the time at which a job is first released in \((\hat{t}, \hat{t})\), and we call the half-open interval \([t_{bi+1}, t_i)\) the principal period and denote it by \(T_{i+1}\). See Figure 1. Then a principal period \(T_k\) is given for some \(k \geq 0\), such

Fig. 1. The division of time intervals.
Lemma 1. In each principal period $T_i$, there is no idle time on any machine in EXF’s schedule between the barrier time $t^b_i$ and the critical time $t_i$, and each job scheduled in $[t^b_i, t_i]$ should have an expiration time $\leq t_i$.

Proof. Fix a principal period $T_i$. Let $I_i$ be the job which is rejected by EXF at $t_i$ and $s_i$ be the release time of $I_i$. From the definition of the barrier time $t^b_i$, we see that no job with expiration time $> t_i$ is scheduled in $[t^b_i, s_i]$ and no idle time exists on any machine in that interval. After $I_i$ is released at $s_i$, no idle time appears on any machine, and EXF never schedules any job with expiration time $> t_i$ until $I_i$ is rejected at $t_i$, because otherwise, EXF could have scheduled $I_i$.

In each principal period $T_i$, we call jobs in $\text{EXF}((t_i, t^b_{i-1}))$ (OPT((t_i, t^b_{i-1}))) late jobs of EXF (OPT, resp.) and $\mathcal{L}_{\text{EXF}}(T_i)$ ($\mathcal{L}_{\text{OPT}}(T_i)$) represents the set of all late jobs of EXF (OPT, resp.) in $T_i$. In the next lemma, it is shown that each late job of OPT should be scheduled by EXF and it is also a late job of EXF in some principal period.

Lemma 2. In each principal period $T_i$, all late jobs of OPT in $T_i$ must be scheduled by EXF and in particular, they are late jobs of EXF in some $T_j$.

Proof. Fix a principal period $T_i$ and let $J$ be a late job of OPT in $T_i$. Then it has the expiration time $e > t_i$. Assume it was not scheduled by EXF before or at $t_i$. Then either, at $t_i$, $J$ was in the queue of EXF or $J$ is released after $t_i$. If $t_i < e < t^b_{i-1}$, $J$ is scheduled in $(t_i, t^b_{i-1})$, because EXF never rejects any job in the queue within the interval. Otherwise, $i.e.$, $e \in T_j$, for some $0 \leq j \leq i - 1$, then in the case the barrier time $t^b_j$ is determined by an idle time, $J$ is scheduled before $t^b_j$, because if it were not scheduled until $t^b_j$, it would be alive in the queue at $t^b_j$ and it would be scheduled. Suppose $t^b_j$ is determined by a job $\hat{J}$ with an expiration time $> t_j$ and $J$ is not scheduled before $t^b_j$. If $t^b_j \leq e \leq t_j$, it is a contradiction, because $\hat{J}$ has the expiration time greater than $J$ but $J$ would be in the queue at the time when $\hat{J}$ was scheduled. So $t_j < e < t^b_{j-1}$. From Lemma 1, $J$ is not scheduled within the interval $[t^b_j, t_j]$ and it is in queue at $t_j$. Since EXF never rejects any job in the queue within $(t_j, t^b_{j-1})$, the job $J$ is scheduled.

If a job $J$ with a processing time $\ell$ starts to process at $s$ and $s \leq t < s + \ell$, the job $J$ is said to cross the time $t$. For each principal period $T_i$, we define frontier jobs of $T_i$ in EXF (OPT) as the jobs scheduled by EXF (OPT, resp.) to cross the time $t_i - [t_i - t^b_i]$, denoted by $f_i$. In particular, $\mathcal{F}(T)$ represents the set of all frontier jobs of a principal period $T$ in EXF. Here we define new periods $\{\hat{T}_i\}$
from the principal periods \( \{ T_i \} \). In nondecreasing order, two adjacent principal periods may be merged into one satisfying the following criterion. Assume new periods \( \tilde{T}_k, k = 0, \ldots, j - 1 \), are defined from \( T_k, k = 0, \ldots, i - 1 \). Let \( t_k \) be the time \( t_k^b + \lfloor t_k - t_k^b \rfloor + 1 \). First set \( \tilde{T}_j \) to be \( T_i \). If \( g_{i+1} > f_i \), then set \( \tilde{T}_j \) to be \( \tilde{T}_j \cup T_{i+1} \), that is, two principal periods \( T_i \) and \( T_{i+1} \) are merged, and otherwise, set \( \tilde{T}_{j+1} \) to be \( T_{i+1} \). Thus the new periods \( \{ \tilde{T}_i \} \) are obtained, each of which consists of either one principal period or two principal periods. We will refer to the sole period as the period which consists of one principal period and the combined period as the period into which two principal periods are merged.

In the following lemmas, we will bound \( |\text{EXF}(\tilde{T}_i)| \) and \( |\text{OPT}(\tilde{T}_i)| \) for each \( i \).

Lemma 3. For each sole period \( \tilde{T}_i \),

\[
|\text{EXF}(\tilde{T}_i)| \geq |\mathcal{F}(T_{h_i}) \cap \text{EXF}(\tilde{T}_i)| + m[t_{h_i} - t_{h_i}^b] + |\mathcal{L}_{\text{EXF}}(T_{h_i})|,
\]

where \( \tilde{T}_i \) is the principal period \( T_{h_i} \).

Proof. Fix a sole period \( \tilde{T}_i \). From Lemma 4, there is no idle time on any machine in \([ t_{h_i}^b, t_{h_i} \] in the schedule of EXF. So EXF can schedule at least \( m[t_{h_i} - t_{h_i}^b] \) jobs which are not the frontier jobs of \( T_{h_i} \) within this interval.

Lemma 4. For each sole period \( \tilde{T}_i \),

\[
|\text{OPT}(\tilde{T}_i)| \leq m([t_{h_i} - t_{h_i}^b] + 1) + |\mathcal{L}_{\text{OPT}}(T_{h_i})|,
\]

where \( \tilde{T}_i \) is the principal period \( T_{h_i} \).

Lemma 4 is obvious, since any scheduler can allocate at most \( m([t_{h_i} - t_{h_i}^b] + 1) \) jobs in \([ t_{h_i}^b, t_{h_i} \], and Lemma 5 can be shown similarly to Lemma 3.

Lemma 5. For each combined period \( \tilde{T}_i \) which consists of two principal periods \( T_{h_i} \) and \( T_{h_{i+1}} \),

\[
|\text{EXF}(\tilde{T}_i)| \geq |\mathcal{F}(T_{h_{i+1}}) \cap \text{EXF}(\tilde{T}_i)| + m([t_{h_i} - t_{h_i}^b] + t_{h_{i+1}} - t_{h_{i+1}}^b) + \sum_{j=h_i}^{h_{i+1}} |\mathcal{L}_{\text{EXF}}(T_j)|.
\]

Lemma 6. For each combined period \( \tilde{T}_i \) which consists of two principal periods \( T_{h_i} \) and \( T_{h_{i+1}} \),

\[
|\text{OPT}(\tilde{T}_i)| \leq m([t_{h_i} - t_{h_i}^b] + [t_{h_{i+1}} - t_{h_{i+1}}^b] + 1) + |\mathcal{L}_{\text{OPT}}(T_{h_i})|.
\]

Proof. Consider frontier jobs of \( T_{h_i} \) in OPT. Then there are at most \( m[t_{h_i} - t_{h_i}^b] \) jobs which are not the frontier jobs of \( T_{h_i} \) and are scheduled by OPT in \([ t_{h_i}^b, t_{h_i} \]. Also since \( g_{h_{i+1}} > f_{h_i} \), at most \( m[t_{h_{i+1}} - t_{h_{i+1}}^b] \) jobs which are not the frontier jobs of \( T_{h_i} \) can be scheduled by OPT in \([ t_{h_{i+1}}^b, t_{h_{i+1}}^b \). Thus since there can be at most \( m \) frontier jobs of \( T_{h_i} \), the upper bound is obtained.
From Lemma 2 we will show the performance of EXF when $|\kappa| \geq 1$.

**Theorem 1.** The EXF algorithm is $(1 + \frac{1}{|\kappa|+1})$-competitive, where $|\kappa| \geq 1$.

**Proof.** From Lemma 2 all late jobs of OPT also are late jobs of EXF, that is, for each sole or combined period $\tilde{T}_i$, $i = 0, \cdots, K$,

$$\sum \left| \mathcal{L}_{opt}(T_{h_i}) \right| \leq \sum \left| \mathcal{L}_{ext}(T_j) \right|,$$

(1)

where the sole or combined period $\tilde{T}_i$ consists of the principal period $T_{h_i}$ or the principal periods $T_{h_i}$ and $T_{h_i+1}$, respectively, and the right sum is done over all principal periods $T_j$. For convenience, assume that $\tilde{T}_K$ is the combined period and there is neither an idle time on some machine nor a job with an expiration time $> t_{h_{K+1}}$ scheduled in $[0, t_{h_{K+1}}]$. Then $|\mathcal{F}(T_{h_{K+1}}) \cap EXF(\tilde{T}_K)| = m$. Thus we obtain the following from Equation (1) and Lemma 3:

$$OPT(J) \leq EXF(J) + Km - \sum_{i=0}^{K-1} |\mathcal{F}(T_{k_i}) \cap EXF(\tilde{T}_i)|,$$

where we define $k_i = h_i$ and $k_i = h_i+1$ for each sole period and combined period $\tilde{T}_i$, respectively. Also the lower bound of $EXF(J)$ is given by the following claim.

**Claim.** $EXF(J) \geq Km[\kappa] + Km - \sum_{i=0}^{K-1} |\mathcal{F}(T_{k_i}) \cap EXF(\tilde{T}_i)|$.

Let $\mathcal{A}_i = \mathcal{F}(T_{k_i}) - EXF(\tilde{T}_i)$ and $\mathcal{B}_i = EXF([t_{k_i}^b, f_{h_i}]) \cup \mathcal{L}_{EXF}(T_{h_i})$. Then $\mathcal{A}_i$ consists of the jobs which are frontier jobs of $T_{k_i}$ and not scheduled within $T_{k_i}$, and if $\tilde{T}_i$ is a combined period, $\mathcal{B}_i$ contains all jobs in $EXF(T_{k_i})$, all frontier jobs of $T_{h_i}$, and all late jobs of EXF in $T_{h_i}$ and if $\tilde{T}_i$ is a sole period, it contains the frontier jobs of $T_{h_i}$ belonging to $EXF(T_{h_i})$ and all late jobs of EXF in $T_{h_i}$.

First we can see from Lemma 1 that for each $i = 0, \cdots, K-1$, $|EXF(\tilde{T}_i) - B_i| \geq m[t_{h_i} - t_{h_i}^b]$. Let $\mathcal{A} = \cup_{i=0}^{K-1} \mathcal{A}_i$ and $\mathcal{B} = \cup_{i=0}^{K} \mathcal{B}_i$. We define a mapping $\mathcal{I} : \mathcal{A} \to \mathcal{B}$ as follows. Given $J \in \mathcal{A}$. Then $J \in \mathcal{F}(T_{k_j}) - EXF(\tilde{T}_j)$, for some $j$. If $J$ is the late job of EXF in $T_{h_{j+1}}$, that is, $J \in EXF([t_{h_{j+1}}^b, t_{k_j}^b])$, then set $\mathcal{I}(J) = J$, and if $J \in EXF([t_{k_j}^b, t_{h_{j+1}}])$, then in the case that $\tilde{T}_j$ is the combined period, there is a
job $J' \in \text{EXF}([t^b_{k_j}, t_{k_j}])$ on the same machine as $J$, because $|t_{k_j} - t^b_{k_j}| \geq \lfloor \kappa \rfloor \geq 1$.

For convenience, we choose the job $J'$ scheduled next to $J$ and set $\mathcal{I}(J) = J'$. In the case that $\tilde{T}_j$ is the sole period, there is a frontier job $J' \in \text{EXF}([t^b_{h_{j+1}}, f_{h_{j+1}}])$ on the same machine as $J$, because otherwise, if the last job in $\text{EXF}([t^b_{h_{j+1}}, f_{h_{j+1}}])$ on the machine is completed at $\tilde{i}$, $\tilde{i} \leq t_{h_{j+1}} \leq f_{k_j}$ since $\tilde{T}_j$ is the sole period, that is, $J$ cannot be the frontier job of $T_k_j$ and it is a contradiction (see Figure 2).

Then we set $\mathcal{I}(J) = J'$. It is obvious that the mapping $\mathcal{I}$ is one to one, which says $|A| \leq |B|$. Let $\mathcal{E}_i = \mathcal{F}(T_{k_i}) \cap \text{EXF}(\tilde{T}_i)$. Then since $|A| = Km - \sum_{i=0}^{K-1} |\mathcal{E}_i|$, we can prove the claim from the following:

\[
\text{EXF}(\mathcal{J}) \geq \sum_{i=0}^{K-1} |\text{EXF}(\tilde{T}_i) - B_i| + |B| \\
\geq m \sum_{i=0}^{K-1} |t_{h_i} - t^b_{h_i}| + |A| \\
\geq Km|\kappa| + Km - \sum_{i=0}^{K-1} |\mathcal{E}_i|.
\]

Consequently, we obtain the desired bound:

\[
\frac{\text{OPT}(\mathcal{J})}{\text{EXF}(\mathcal{J})} \leq 1 + \frac{Km - \sum_i |\mathcal{E}_i|}{Km|\kappa| + Km - \sum_i |\mathcal{E}_i|} \leq 1 + \frac{1}{\lfloor \kappa \rfloor + 1}.
\]

### 2.2 Case $\lfloor \kappa \rfloor = 0$

In this subsection, when $\lfloor \kappa \rfloor = 0$, it is shown that the EXF algorithm is 2-competitive. We obtain the set of the principal periods $\{T_i\}$ as in the previous and merge them differently. New periods $\tilde{T}_i$ are obtained as follows: Assume $\tilde{T}_k$, $k = 0, \ldots, j - 1$, are obtained from $T_k$, $k = 0, \ldots, i - 1$. If there is a job which starts to process in $[t^b_{l}, t_{l}]$ and is completed after $t_i$, for some $l > i$, then let $T_u$ be the principal period having the largest index $u > i$ in which such a job exists, and we merge the principal periods $T_i, \ldots, T_u$ into one as $\tilde{T}_j$. Otherwise, set $\tilde{T}_j$ to be $T_i$. Thus the periods $\tilde{T}_i$ which consist of one principal period and principal periods more than one are called long periods and short periods, respectively. Let $\mathcal{J}_i$ and $\tilde{\mathcal{J}}_i$ be the set of all jobs which is released within $T_i$ and $\tilde{T}_i$, respectively. Then we compare $\text{EXF}(\mathcal{J}_i)$ with $\text{OPT}(\tilde{\mathcal{J}}_i)$ for each $i$. In each $\mathcal{J}_i$, the jobs having expiration times $> t_i$ are called patient jobs and it can be shown similarly to Lemma 2 that all patient jobs should be scheduled by EXF. Also we note that in each $T_i$, all jobs scheduled by EXF in $[t^b_{i}, t_{i}]$ belong to $\mathcal{J}_i$.

**Lemma 7.** In each principal period $T_i$, all jobs scheduled by EXF in $[t^b_{i}, t_{i}]$ belong to $\mathcal{J}_i$, that is, they are released after or at the barrier time $t^b_{i}$ of $T_i$, and all patient jobs of $\mathcal{J}_i$ must be scheduled by EXF.
Lemma 8. For each long period $\tilde{T}_i$,

$$OPT(\tilde{J}_i) \leq EXF(\tilde{J}_i) + m.$$ 

Lemma 9. For each short period $T_i$,

$$OPT(\tilde{J}_i) \leq EXF(\tilde{J}_i) + m.$$ 

Proof. Fix a short period $T_i$ which consists of principal periods $T_j, \cdots, T_k$. If $g_k = t_k^b + [t_k - t_k^b] + 1 \leq t_j$, then since $T_i$ is the short period, we see that $t_j - t_k < 1$. Thus $OPT$ can schedule at most $m([t_k - t_k^b] + 2)$ jobs in $[t_k^b, t_j]$. So the upper bound of $OPT(\tilde{J}_i)$ is given by

$$OPT(\tilde{J}_i) \leq m([t_k - t_k^b] + 2) + |P_j|,$$

where $P_j$ is the set of all patient jobs of $J_j$. But $EXF$ schedules at least $m[t_k - t_k^b]$ jobs which are not the frontier jobs of $T_k$ in $[t_k^b, t_k]$. Also the $m$ frontier jobs of $T_j$ exist, and if some of them start to process before or at $t_k$, there is the frontier job of $T_k$ on the same machine, because $g_k \leq t_j$. Thus we can obtain the $m$ jobs of $\tilde{J}_i$ each of which is either a frontier job of $T_j$ scheduled after $t_k$ or a frontier job of $T_k$. Therefore the following is given from Lemma 7

$$EXF(\tilde{J}_i) \geq m([t_k - t_k^b] + 1) + |P_j|.$$

If $g_k > t_j$, then $OPT$ can schedule at most $m([t_k - t_k^b] + 1)$ jobs in $[t_k^b, t_k]$ but $EXF$ schedules at least $m[t_k - t_k^b]$ jobs in $[t_k^b, t_k]$. Thus the desired equation is obtained.

Theorem 2. The EXF algorithm is 2-competitive, where $|\kappa| = 0$. 

Proof. Given the long or short periods $\tilde{T}_i, i = 0, \cdots, K - 1$. Actually it remains to prove $\sum_i EXF(\tilde{J}_i) \geq Km$. Fix $\tilde{T}_i$ and we assume that $\tilde{T}_i$ consists of the principal periods $T_j, \cdots, T_{j+k}$, where $k \geq 0$, and $k = 0$ if $\tilde{T}_i$ is the long period. Then from the definition of the critical time $t_j$ we can see that there are exactly $m$ jobs crossing $t_j$. Thus we assign the $m$ jobs to $\tilde{T}_i$. If $\tilde{T}_i$ is the long period, any job scheduled before or at $t_{j+1}$ cannot be completed after $t_j$, and if $\tilde{T}_i$ is the short period, any job scheduled before or at $t_{j+k+1}$ cannot be completed after $t_j$, because if there is such a job, it is a contradiction to the maximality of $T_{j+k}$ in the definition of $\tilde{T}_i$. Thus we can see that the $m$ jobs assigned to $\tilde{T}_i$ cannot be re-assigned by any other periods and the desired fact is proved. Consequently, it is followed from Lemma 8 and 9 that

$$OPT(\mathcal{J}) \leq EXF(\mathcal{J}) + Km \leq 2EXF(\mathcal{J}).$$

Until now, we have proved that EXF is $(1 + \frac{1}{|\kappa|+1})$-competitive, for $\kappa \geq 0$. Next, it is shown that the competitive ratio is tight by giving the lower bound. The instance of the lower bound is given just by making $m$ copies for each job in the instance given in [8].
Lemma 10. For $\kappa \geq 0$,

$$\max_{\mathcal{J}} \frac{\text{OPT}(\mathcal{J})}{\text{EXF}(\mathcal{J})} \geq 1 + \frac{1}{\lfloor \kappa \rfloor + 1}$$

and, for any online algorithm $A$,

$$\max_{\mathcal{J}} \frac{\text{OPT}(\mathcal{J})}{A(\mathcal{J})} \geq 1 + \frac{1}{2\lfloor \kappa \rfloor + 3}.$$ 

3 Arbitrary Processing Time Jobs

As in the last section, the principal periods $T_i$ are defined, and the barrier times $t^b_i$ and the critical times $t_i$ which determine $T_i$ are given. Let $\mathcal{J}_i$ be the set of all jobs which are released within $T_i$. It can be shown similarly to Lemma 7 that all patient jobs of $\mathcal{J}_i$ having expiration times $> t_i$ should be scheduled by EXF.

Lemma 11. In each principal period $T_i$, there is no idle time on any machine between the barrier time $t^b_i$ and the critical time $t_i$, and all patient jobs of $\mathcal{J}_i$ must be scheduled by EXF. Thus we get

$$\text{EXF}(\mathcal{J}) \geq \max \{ \sum_i m[t_i - t^b_i], \sum_i \sum_{j \in \mathcal{P}_i} p_j \},$$

where $\mathcal{P}_i$ is the set of all patient jobs of $\mathcal{J}_i$.

Lemma 12. For each principal period $T_i$,

$$\text{OPT}(\mathcal{J}_i) \leq (1 + \frac{1}{\kappa})m[t_i - t^b_i] + \sum_{j \in \mathcal{P}_i} p_j,$$

where $\mathcal{P}_i$ is the set of all patient jobs of $\mathcal{J}_i$.

Proof. Let $J$ be a job of $\mathcal{J}_i$ scheduled to cross the time $t_i$ by OPT. In fact, there can be at most $m$ jobs crossing $t_i$. If the job $J$ is not the patient job of $\mathcal{J}_i$, its slack time is no more than $|t_i - t^b_i|$. From the slackness condition, the processing time $\ell$ of $J$ is less than or equal to the slack time divided by $\kappa$, that is, $\ell \leq \frac{|t_i - t^b_i|}{\kappa}$. Furthermore, the total gain of all the jobs in $\mathcal{J}_i$ which can be scheduled to be completed before $t_i$ is at most $m[t_i - t^b_i]$. Thus we obtain the bound.

Now, we prove that EXF is $(2 + \frac{1}{\kappa})$-competitive, for $\kappa > 0$, and we can also obtain the instance of the lower bound similarly to in [8].

Theorem 3. The EXF algorithm is $(2 + \frac{1}{\kappa})$-competitive, where $\kappa > 0$. 
Proof. From Lemma 11 and 12, it is shown that

\[ OPT(J) \leq \sum_i (1 + \frac{1}{\kappa}) m[t_i - t^b_i] + \sum_i \sum_j p_j \]

\[ \leq (1 + \frac{1}{\kappa}) \text{EXF}(J) + \text{EXF}(J) = (2 + \frac{1}{\kappa}) \text{EXF}(J). \]

Lemma 13. For \( \kappa > 0 \),

\[ \max_J \frac{OPT(J)}{\text{EXF}(J)} \geq 2 + \frac{1}{\kappa}. \]

References

Competitive Online Scheduling with Level of Service⋆
(Extended Abstract)

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Abstract. Motivated by an application in thinwire visualization, we study an abstract on-line scheduling problem. Unlike most scheduling problems, our schedulers can gain partial merit from a partially served request. Thus our problem embodies a notion of “Level of Service” that is increasingly important in multimedia applications. We give two schedulers FirstFit and EndFit based on two simple heuristics, and generalize them into a class of greedy schedulers. We show that both FirstFit and EndFit are 2-competitive, and any greedy scheduler is 3-competitive. These bounds are shown to be tight.

1 Introduction

We study an abstract on-line scheduling problem motivated by visualization across a “thinwire” network [4, 3]. An example of such a visualization problem is a server-client model where the server and client are connected by a thinwire (that is, a bandwidth-limited connection such as the Internet), with the server holding a very large image that the client wishes to visualize. The viewer on the client side can control the transmission process by moving a mouse cursor over a low-resolution copy of the image to be visualized. This mouse motion generates, in real-time, a sequence of sampled positions along the mouse cursor trajectory. Each sampled position \((x, y)\) corresponds to a request for higher resolution data at the position. As the bandwidth is limited, we could only partially serve each request. This is where an on-line scheduler is needed to optimize the decisions. In most scheduling problems, a partially served request does not contribute to the performance of the scheduler. However, in this problem, a partially sent data can still provide useful information to the user. Thus, instead of sending all the requested data, the server has the option of lowering the “level” of the requested service in order to gain an overall better response time. This paper focuses on this level of service property. Note that there is considerable interest in similar Quality of Service (QoS) issues in multimedia research.

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We use the standard notion of “competitiveness” in the sense of Sleator and Tarjan [8], to judge the quality of our online schedulers. A scheduler $S$ produces a feasible schedule $S(I)$ for each instance $I$ of our scheduling problem. Each $S(I)$ has an associated merit, where $\text{merit}(S(I)) \geq 0$. Let $\text{opt}(I)$ denote any feasible schedule for $I$ that maximizes the merit. We say $S$ is $c$-competitive ($c \geq 1$) if for all $I$,

$$\text{merit}(\text{opt}(I)) \leq c \cdot \text{merit}(S(I)) + b,$$

where $b$ is a fixed constant. The competitive ratio of $S$ is defined by

$$C(S) := \sup_I \frac{\text{merit}(\text{opt}(I))}{\text{merit}(S(I))}.$$ 

Thus, we want schedulers $S$ with $C(S) \geq 1$ as small as possible. There is a fairly large literature on competitive algorithms (e.g., [1, 2, 7]). The class of problems most closely related to ours is the online interval packing problem for a single server, where a schedule is a subset of non-overlapping intervals. Lipton and Tomkins [6] study a variant where the input intervals are sorted by their left endpoints. They give a randomized scheduler that is 2-competitive. As we will see, our problem is different from theirs in several ways. Woeginger [9] studied a problem that has several of the features of our problem. Other online interval packing problems can be found in [10, 5].

2 Problem Formulation

We formalize our problem as an on-line scheduling problem. Each request $q$ has four parameters

$$q = (s, t, v, w),$$

where $s$ the start time, $t$ the termination time (or deadline), $v$ is the volume (or size), and $w$ is the weight. We require

$$v \geq 0 \quad \text{and} \quad w \geq 0.$$

Write $\text{st}(q)$, $\text{dl}(q)$, $\text{sz}(q)$, $\text{wt}(q)$ for the above parameters of $q$, respectively. Call the half-open interval $(s, t]$ the span of $q$, written $\text{span}(q)$. A request $q$ can only be served within its span $(s, t]$, and at any time moment $t_0$, at most one request can be served.

An instance $I$ is a sequence $\langle q_1, q_2 \ldots q_n \rangle$ of requests where the start times of the $q_i$’s are in increasing order: note that we allow $\text{st}(q_i) = \text{st}(q_{i+1})$ even though we nominally say $q_i$ starts before $q_{i+1}$. How requests are served is described by the schedule. Formally, a schedule for $I$ is a piece-wise constant function

$$H : \mathbb{R} \to \{q_1, q_2, \ldots, q_n\} \cup \{\emptyset\},$$

where $|H^{-1}(q_k)| \leq \text{sz}(q_k)$ and $H^{-1}(q_k) \subseteq \text{span}(q_k)$ for $k = 1, \ldots, n$. Intuitively, $H(t_0) = q_k$ means the $k$th request is served at time $t_0$ and $H(t_0) = \emptyset$ means
no request is being served. A time moment \( t_0 \) is called a **breakpoint** if \( H \) is discontinuous at \( t_0 \). (More precisely, for every \( \varepsilon > 0 \), there exists \( \delta_i \) \((0 < \delta_i < \varepsilon, i = 1, 2)\) such that \( H(t_0) = H(t_0 - \delta_1) \neq H(t_0 + \delta_2) \)). We further require a schedule \( H \) to have finitely many breakpoints. In addition, for each request \( q \), \(|H^{-1}(q)| \leq \text{sz}(q)\). A half-open interval of the form \((t_0, t_1]\) is called a **time-slot**. Without loss of generality, we may assume \( H^{-1}(q) \) is a finite union of time slots.

The merit \( \text{merit}(H) \) of a schedule is

\[
\sum_{j=1}^{n} \text{wt}(q_j)|H^{-1}(q_j)|.
\]

Relative to a schedule \( H \) at any time \( t_0 \), we call

\[
v' := \text{sz}(q) - |H^{-1}(q) \cap (-\infty, t_0]|
\]

the residual size of \( q \). A request \( q \) is completely served if \( v' = 0 \). If \( v' > 0 \) and \( \text{st}(q) \leq t_0 \leq \text{dl}(q) \), then we say \( q \) is pending. The residue of a pending \( q \) at time \( t_0 \) is the modified request \( q' = (t_0, \text{dl}(q), v', \text{wt}(q)) \).

For each completely served request, the scheduler gains \( \text{sz}(q)\text{wt}(q) \) merit points (so weights are multiplicative). Moreover, partially served requests gain a proportional fraction of this merit. This is unlike usual scheduling problems in which partially served requests receive no merit. Our model is meaningful for the “foveated visualization” studied in [4, 3] because a scheduler can reduce the amount of requested visualization data along a finely graduated scale. This can be achieved by reducing two parameters, the **foveal radius** and/or **foveal resolution**. The foveal radius measures how fast the resolution falls away from the foveal center, and the foveal resolution measures the maximum resolution (which is at the foveal center).

**Preemption.** It is implicit in the above definitions that the servicing of any request can be preempted as often as we like with no penalty. Hence we may imagine the scheduler to “plan” a schedule based on all the currently residual requests. It services the requests according to this plan until the arrival of a new request. Then it suspends the current plan, recomputes a new plan based on the new set of residual requests, and repeats the process.

**Optimal Schedules.** We say \( H \) is **optimal** for \( I \) if \( \text{merit}(H) \) is maximum among all schedules for \( I \). The existence of optimal schedules is not immediate.

**Lemma 1.** For all sequences \( I \) of \( n \) requests, there exists an optimal schedule with at most \( 2n^2 + n \) breakpoints.

The optimal schedules for an instance may not be unique. In the full paper, we give a canonical representation whereby the optimal schedule is unique. Let \( \text{opt}(I) \) denote the canonical optimal schedule for \( I \).

**Ordering of Requests.** The schedulers in this paper make decisions by giving priority to heavier weighted requests. In case \( \text{wt}(p) = \text{wt}(q) \), we resolve the tie by treating \( p \) as “heavier” than \( q \) if and only if \( p \) starts before \( q \).
2.1 FirstFit

The online scheduler that always serves the heaviest residual request at each moment is called the FirstFit scheduler. Figure 1 shows the schedule produced by FirstFit on an instance of two requests $q_1$ and $q_2$. Although this example may appear contrived, we can modify $q_1$ to $\tilde{q}_1$ where $\tilde{q}_1 = (0, 2, 1, 1 + \epsilon)$. For any $\epsilon > 0$, the FirstFit schedule is the one shown in Figure 1.

![Diagram](image)

**Fig. 1.** The top figure illustrates the instance $I = \langle q_1, q_2 \rangle$ where $q_1 = (0, 2, 1, 1)$ and $q_2 = (0, 1, 1, 1)$. Each horizontal “dashed” line represents the span of the request. Although $wt(q_1) = wt(q_2)$, $q_1$ is “heavier” than $q_2$ by our tie-breaking rule. In $\text{opt}(I)$, $q_1$ and $q_2$ are served in the time-slots $[1, 2]$ and $[0, 1]$, respectively. However, in $\text{FirstFit}(I)$, only $q_1$ is served.

2.2 EndFit

Consider an online scheduler which always serves according to the optimal schedule for the current set of residual requests. This was first suggested by Estie Arkin\(^1\). To implement such a scheduler, we can invoke a general off-line algorithm for computing optimal schedules upon each new arrival of a request. But it turns out that a very simple scheduler can be used. This scheduler, on an arbitrary instance $I$, operates as follows:

Starting from the heaviest request down to the lightest, allocates each request $q \in I$ to the latest available time-slot(s) within $\text{span}(q)$.

Call this the OffEndFit scheduler. It is an off-line algorithm because it must see the entire set of requests to make its decisions. This scheduler is optimal for a special class of instances.

**Lemma 2.** If $I$ is an instance in which all requests have a common starting time, then $\text{OffEndFit}(I)$ is the canonical optimal schedule for $I$.

Let EndFit be the online scheduler which always serves according to the OffEndFit schedule for the residual requests. More precisely, on arrival of a new

\(^1\) Private communication (1997).
request \( q \), \text{EndFit} preempts the current service. It computes a new schedule \( P \) for the current residual requests using \text{OffEndFit}, and continues by servicing \( P \). Call \( P \) the \textit{plan} upon arrival of \( q \). Since all residual requests have a common starting time, \( P \) is the canonical optimal schedule. Figure 2 shows the \text{EndFit} schedule for an instance \( I = (q_1, q_2) \). The \text{EndFit} schedule for this example may appear contrived. To see that it is “correct in the limit”, let \( I_{\epsilon} = (q_0, q_1, q_2) \) where \( q_0 = (0, 1, 1, \epsilon) \). For any \( 0 < \epsilon < 1 \), the off-line optimal schedule for the current residual requests is unique. As \( \epsilon \to 0 \), \text{EndFit}(I_\epsilon) approaches the one schedule shown in Figure 2.

![Fig. 2.](image)

**Fig. 2.** The top figure illustrates the instance \( I \) of two requests: \( q_1 = (0, 2, 1, 1) \) and \( q_2 = (1, 2, 1, 1) \). In the \text{opt}(I), \( q_1 \) and \( q_2 \) are served. However, in \text{EndFit}(I), only \( q_1 \) is served.

### 3 Competitive Ratio of FirstFit

Example 1 (Figure 1) shows that the competitive ratio of \text{FirstFit} is at least 2. We will show that \text{FirstFit} is 2-competitive. Before presenting the proof, let us give two definitions.

**Charging Scheme.** Let \( H, H_1 \) and \( H_2 \) be schedules for an instance \( I \). We often need to argue that the merit of \( H \) is no larger than the sum of the merits of \( H_1 \) and \( H_2 \). Our approach is to \textit{charge} a portion of \( H \) to \( H_1 \) and the remaining to \( H_2 \). Intuitively, the charging process can be viewed as first cutting \( H_1 \) and \( H_2 \) into pieces and then piecing them together again to form another piecewise-constant function \( H_{\text{chg}} \). Each piece, after cutting, may be translated before being placed into their slot in \( H_{\text{chg}} \). As it may turn out that \( |H_{\text{chg}}^{-1}(q)| > sz(q) \) for some request \( q \), \( H_{\text{chg}} \) is not necessarily a schedule. The cut-and-paste is done in a way that for all \( t \), \( wt(H_{\text{chg}}(t)) \geq wt(H(t)) \). Therefore,

\[
\text{merit}(H) \leq \text{merit}(H_{\text{chg}}) \leq \text{merit}(H_1) + \text{merit}(H_2).
\]  

In particular, if \( H_1 = H_2 \), then we have \( \text{merit}(H) \leq 2 \cdot \text{merit}(H_1) \). When we use the phrase: “charge \((s', e']\) from \( H \) to \( H_1 \) at \((s, e]\)”, we mean that a piece \((s, e]\) is cut from \( H_1 \) and placed into the slot \((s', e']\) in \( H_{\text{chg}} \). Equivalently, if \( H^{-1}(q) \)
is the interval \((s', e']\), we may say that the request \(q\) is charged from \(H\) to \(H_1\) at \((s, e]\) (see Figure 3). Thus, to show (1), we need to charge each time slot of \(H\) to either \(H_1\) or \(H_2\), and ensure that no part of \(H_1\) or \(H_2\) is charged more than one.

![Fig. 3. Charging the request \(q\) in \(H\) to \(H_1\) at \((s, e]\).](image)

**Intactness.** A request \(q\) is *intact* in a schedule \(H\) if \(H^{-1}(q)\) is connected, and either \(|H^{-1}(q)| = 0\) or \(sz(q)\). Most of our proofs will be simplified if we assume intactness.

**Theorem 1.** For any instance \(I\),

\[
\text{merit}(\text{opt}(I)) \leq 2 \cdot \text{merit}(\text{FirstFit}(I)).
\]

**Proof.** Given an instance \(I\), let \(H_{\text{ff}} := \text{FirstFit}(I)\) and \(H_{\text{opt}} := \text{opt}(I)\). We can assume that requests in \(I\) are intact in both \(H_{\text{ff}}\) and \(H_{\text{opt}}\).

Let \(H_0\) be an identical copy of \(H_{\text{ff}}\). We want to charge requests served in \(H_{\text{opt}}\) to \(H_0\) and \(H_{\text{ff}}\). Let \(\{t_1, t_2, \ldots, t_m\}\) be the distinct breakpoints in \(H_{\text{opt}}\), where \(t_i < t_j\) if and only if \(i < j\).

For each \(t_i\), starting from \(i := 1\) to \(m - 1\), consider the time-slot \((t_i, t_{i+1}]\). Let \(q_{\text{opt}} := H_{\text{opt}}(t_{i+1})\). Let \(q_{\text{ff}}\) be the lightest request served during \((t_i, t_{i+1}]\) by \(\text{FirstFit}\). There are two cases:

1. If the request \(q_{\text{ff}}\) is not lighter than \(q_{\text{opt}}\), charge \(q_{\text{opt}}\) from \(H_{\text{opt}}\) to \(H_{\text{ff}}\) at \((t_i, t_{i+1}]\).
2. Otherwise, charge \(q_{\text{opt}}\) from \(H_{\text{opt}}\) to \(H_0\) at \(H_0^{-1}(q_{\text{opt}})\).

We have to show that in the second case, \(|H_0^{-1}(q_{\text{opt}})| \geq |(t_i, t_{i+1}]|\). In the first place, why is the weight of \(q_{\text{ff}}\) lighter? The request \(q_{\text{ff}}\) is chosen by \(\text{FirstFit}\) because it is the heaviest request among the pending requests. This implies that \(q_{\text{opt}}\) is not a pending request, even though \(t_i\) is in the span of \(q_{\text{opt}}\). So \(q_{\text{opt}}\) must have been completely served by \(\text{FirstFit}\). This implies that \(|H_0^{-1}(q_{\text{opt}})| \geq |(t_i, t_{i+1}]|\).

### 4 Competitive Ratio of EndFit

Example 2 (Figure 2) shows that the competitive ratio of \(\text{EndFit}\) is at least 2. We now show that this constant is the best possible.
The upper bound proof is considerably more subtle than the proof for FirstFit. The key result is Theorem 2 below which formalizes this observation about EndFit: it never hurts the performance of EndFit to have a request started at an earlier time. For example, in Figure 2, the performance of EndFit will improve if the request q2 starts at an earlier time. The analogous fails for FirstFit. For example, in Figure 1, the performance of FirstFit would improve if q1 starts at a time later than 0.

A request \( \tilde{q} \) is a trimmed version of q if \( st(\tilde{q}) \geq st(q) \), \( dl(\tilde{q}) = dl(q) \), \( wt(\tilde{q}) = wt(q) \) and \( sz(\tilde{q}) \leq sz(q) \). Thus, a trimmed version of q may start later than the original q. An instance \( \tilde{I} \) is a trimmed instance of I if there is a one-one (not necessarily onto) mapping from \( \tilde{I} \) to I such that any \( \tilde{q} \) in \( \tilde{I} \) is a trimmed version of its corresponding request in I. Clearly, \( merit(\text{opt}(\tilde{I})) \leq merit(\text{opt}(I)) \). Similar relationship also holds for EndFit.

**Theorem 2.** If \( \tilde{I} \) is a trimmed version of I then

\[
\text{merit(EndFit(\tilde{I}))} \leq \text{merit(EndFit(I))}.
\]

We now use Theorem 2 to show that EndFit is 2-competitive.

**Theorem 3.** For any instance I,

\[ \text{merit(\text{opt}(I))} \leq 2 \cdot \text{merit(EndFit(I))}. \]

**Proof.** We can assume that requests in I are intact in \( \text{opt}(I) \). Let \( \tilde{I} \) be the trimmed instance of I such that for any request \( q \in I \), if \( (\text{opt}(I))^{-1}(q) = (t_1, t_2) \), then the corresponding trimmed request \( \tilde{q} \) satisfies \( st(\tilde{q}) := t_1 \) and \( sz(\tilde{q}) := t_2 - t_1 \); otherwise if \( \text{opt}(I)^{-1}(q) = \emptyset \), then \( \tilde{q} \) satisfies \( sz(\tilde{q}) := 0 \). We can further assume that \( \tilde{I} \) is intact in all the plans of EndFit with \( \tilde{I} \). (Recall that a plan is the optimal schedule for the residues). By definition, we have

\[
\text{merit(\text{opt}(I))} = \text{merit(\text{opt}(\tilde{I}))}. \tag{2}
\]

The instance \( \tilde{I} \) has the nice property that requests arrive at a “constant rate”, that is, if a request q starts at time t, then no other request starts during \( (t, t + sz(q)) \). Let \( H_1 \) and \( H_2 \) be two identical copies of EndFit(\( \tilde{I} \)). Our theorem is proved if we show how to charge requests in \( \text{opt}(\tilde{I}) \) to \( H_1 \) and \( H_2 \).

Consider a request q in \( \tilde{I} \). Let \( P^+ \) be the plan upon the arrival of q and \( P^- \) be the plan just before the arrival of q. There are two cases.

1. If q is allocated in the new plan \( P^+ \), then it is possible that there are some requests which are originally allocated in \( P^- \), but not in the new plan. Call these requests the ousted requests.
2. Otherwise, call q the ousted request.

Let s be the total size of the ousted requests. Note that \( s \leq sz(q) \) and the total merit of the ousted requests is not more than the total merit of the requests
allocated in \((\text{st}(q), \text{st}(q) + s]\) in \(P^+\). Furthermore, the new plan will be carried out without interruption at least until \(\text{st}(q) + \text{sz}(q)\). Charge the ousted requests to \(H_2\) at \((\text{st}(q), \text{st}(q) + \text{sz}(q)]\) and the served requests during \((\text{st}(q), \text{st}(q) + \text{sz}(q)]\) to \(H_1\) at \((\text{st}(q), \text{st}(q) + \text{sz}(q)]\). The above is a valid charging scheme. Thus, we have

\[2 \cdot \text{merit}(\text{EndFit}({\tilde{I}})) \geq \text{merit}(\text{opt}({\tilde{I}})).\]

By Theorem 2 and (2), we have

\[2 \cdot \text{merit}(\text{EndFit}(I)) \geq \text{merit}(\text{opt}(I)).\]

5 A Class of Greedy Schedulers

Looking at the behavior of \(\text{EndFit}\) and \(\text{FirstFit}\) on specific examples, it appears that they are complementary in the sense that if \(\text{EndFit}\) performs poorly on an instance, then \(\text{FirstFit}\) will perform well, and vice-versa. This suggests studying some combination of these two heuristics and motivates the generalization to a class of Greedy schedulers.

A scheduler \(S\) in this class behaves as follows.

(A) At the moment a new request \(q\) starts, it suspends the current service (that is, it preempts the currently served request).

(B) Scheduler \(S\) computes a new plan \(H\), which is a schedule for the set of residues of currently pending requests. We call \(H\) a ‘plan’ because the scheduler may not carry out the schedule as planned due to the subsequent new requests. The plan \(H\) is computed by considering the residues one by one, starting from the heaviest request down to the lightest request. Let \(p\) be the request being considered and call \(|H^{-1}(p)|\) the allocation to \(p\). The allocation to \(p\) is subjected to the following restriction:

(*) The allocation to \(p\) must be maximized. For example, if it is possible to completely allocate \(p\), the whole of \(p\) must be allocated. However, there is no restriction on where \(p\) is allocated. Time-slots, once allocated, are not subsequently revised in creating this plan.

(C) It carries out the plan until a new request starts, whereupon we go back to step (A).

Different members of the Greedy class differ only in their strategies for (B) subjected to the restriction(*). Note that our first example \(\text{FirstFit}\) is a Greedy scheduler: the request \(p\) in (*) is allocated in the earliest possible time-slots. The second example \(\text{EndFit}\) is also a greedy scheduler. Its strategy for (B) is rather counter-intuitive: the request \(p\) is allocated in the latest possible time slots.

By combining the counter examples for \(\text{FirstFit}\) and \(\text{EndFit}\), we can find a Greedy scheduler whose competitive ratio \(\geq 3\). The next theorem shows that this bound of 3 cannot be improved.

**Theorem 4.** Every Greedy scheduler is 3-competitive.
6 General Lower Bound

From the previous section, we know that every greedy scheduler is 3-competitive. Are there schedulers outside the Greedy class with competitive ratio less than 2? We note a partial result in this direction: every deterministic scheduler has competitive ratio at least \(2(2 - \sqrt{2}) > 1.17\).

In proof, consider this adversary: at time 0, the adversary releases two requests \(q_0 := (0, 2, 1, 1)\) and \(q_1 := (0, 1, 1, \sqrt{2} - 1)\). At time \(t = 1\), let the residual size of \(q_0\) be \(s_0\). If \(s_0\) is less than \(1/2\) then the request \(q_2 := (1, 2, 1, 1)\) is released. Otherwise, no further requests will be released. It may be verified that any deterministic scheduler achieve a merit of at most \(\frac{1}{2(2 - \sqrt{2})}\) of the maximum possible.

Unfortunately, this lower bound of 1.17 leaves a wide gap from the current upper bound of 2. On the other hand, no simple variation of this adversary seems to give a better lower bound.

7 On the Number of Breakpoints and Multi-tasking

Both FirstFit and EndFit make \(O(n)\) breakpoints where \(n\) is the number of requests. Does the number of breakpoints affect the performance of a scheduler? We give an alternative formulation of the scheduling problem which could be viewed as allowing an infinite number breakpoints. In this multi-tasking environment, several requests can be served concurrently, but each at a (possibly) different rate. However, in any time interval of size \(\Delta t\), the total size of requests served within this interval must not exceed \(\Delta t\). Equivalently, we allow “fractional service” where the total service at any moment sums to 1.

A concrete example of such a scheduler is FirstEndFit: It simulates FirstFit and EndFit concurrently and serves half of what FirstFit and EndFit would serve. That is, if FirstFit and EndFit will serve \(q\) and \(p\) respectively in the time slot \((s_0, t_0]\), then FirstEndFit will serve \(p\) and \(q\) concurrently but each at half the rate. We suspect that FirstEndFit is \((3/2)\)-competitive.

Note that FirstEndFit can be also viewed as the following randomized scheduler under the original single-tasking setting: Before receiving any request, it tosses a fair coin; If the outcome is head, then it simulates FirstFit, otherwise it simulates EndFit.

Clearly, the expected merit gained by this randomized scheduler is same as the merit gained by FirstEndFit.

8 Conclusion

We have formulated a “level-of-service” scheduling problem that arises naturally in our thinwire visualization applications. This formulation is also useful in real-time systems where quality of jobs can be traded-off for time. We have derived several competitive algorithms in this setting. We continue to investigate the
many interesting questions that are open. Besides sharpening the results in the
paper, we pose the following directions for further work: (1) Find optimal sched-
ulers which are not restricted to be on-line. (2) Study the problem with other
measures of merit (instead of multiplicative weights in this paper). (3) Introduce
a model of penalty for preemption. (4) Introduce randomization.

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On-Line Variable Sized Covering

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Abstract. We consider one-dimensional and multi-dimensional vector covering with variable sized bins. In the one-dimensional case, we consider variable sized bin covering with bounded item sizes. For every finite set of bins $B$, and upper bound $1/m$ on the size of items for some integer $m$, we define a ratio $r(B, m)$. We prove this is the best possible competitive ratio for the set of bins $B$ and the parameter $m$, by giving both an algorithm with competitive ratio $r(B, m)$, and an upper bound of $r(B, m)$ on the competitive ratio of any on-line deterministic or randomized algorithm. The ratio satisfies $r(B, m) \geq m/(m+1)$, and equals this number if all bins are of size 1.

For multi-dimensional vector covering we consider the case where each bin is a binary d-dimensional vector. It is known that if $B$ contains a single bin which is all 1, then the best competitive ratio is $\Theta(1/d)$. We show an upper bound of $1/2^{d(1-o(1))}$ for the general problem, and consider four special case variants. We show an algorithm with optimal competitive ratio $1/2$ for the model where each bin in $B$ is a unit vector. We consider the model where $B$ consists of all unit prefix vectors. Each bin has $i$ leftmost components of 1, and all other components are 0. We show that this model is harder than the case of unit vector bins, by giving an upper bound of $O(1/\log d)$ on the competitive ratio of any deterministic or randomized algorithm. Next, we discuss the model where $B$ contains all binary vectors. We show this model is easier than the model of one bin type which is all 1, by giving an algorithm of ratio $\Omega(1/\log d)$.

The most interesting multi-dimensional case is $d = 2$. Previous results give a 0.25-competitive algorithm for $B = \{(1, 1)\}$, and an upper bound of 0.4 on the competitive ratio of any algorithm. In this paper we consider all other models for $d = 2$. For unit vectors, we give an optimal algorithm with competitive ratio $1/2$. For unit prefix vectors we give an upper bound of $4/9$ on the competitive ratio of any deterministic or randomized algorithm. For the model where $B$ consists of all binary vectors, we design an algorithm with ratio larger than 0.4. These results show that all above relations between models hold for $d = 2$ as well.

1 Introduction

We consider on-line problems of covering variable sized bins. The one-dimensional version of this problem is defined as follows: We are given a finite set $B$ of allowed bin sizes, each $b \in B$ satisfies $0 < b \leq 1$ and the largest element of $B$ is 1.
Items with size in \((0, 1]\) arrive on-line, each item is to be assigned to a bin upon arrival. The algorithm may assign a new item to a previously used bin, or open a new bin of any size in \(B\). A bin of size \(b\) is covered if the total size of items assigned to it is at least \(b\). The goal of an algorithm is to maximize the sum of the sizes of all covered bins. This sum is the value of an algorithm. In this paper we consider bin covering with bounded item sizes. The size of each arriving item is at most \(1/m\) where \(m\) is a fixed integer.

The multi-dimensional version is defined as follows: For a dimension \(d\) we are given a set of \(d\)-dimensional bins \(B\), which is a subset of \(\{0, 1\}^d\), i.e., all components of the allowed bins are binary. The “vector weight” of a \(d\)-dimensional vector \(b\) is the sum of its components, this applies both to multi-dimensional bins (“bin weights”) and to multi-dimensional items (“item weights”). In the one-dimensional case the “bin weight” of a bin is simply its size and the “item weight” of an item is the size of the item. Vector items with size in \([0, 1]^d\) arrive on-line, and as in the one-dimensional case, each new item is to be assigned to a new or previously used bin upon arrival. A bin of size \(b\) is covered if the vector \(x\), which is the sum of all items assigned to the bin, satisfies \(x \geq b\). The goal of an algorithm is to maximize the sum of the bin weights of all covered bins. This number is the value of the algorithm. We study the general problem of arbitrary sets \(B\), and also consider three interesting cases of multi-dimensional variable sized vector covering. The first case is where \(B\) consists of all possible binary vectors. The second case is when all bins are unit vectors (have only one non-zero coordinate). In the last case we consider a unit prefix sequence of bins, each bin has \(i\) leftmost 1 coordinates, and all others are 0.

**Applications**: Variable sized bin and vector covering relate to a case where a large number of jobs of a finite type are to be done (these are the bins). Each job is done by one worker or by a combined work of more than one worker (the workers are the items). An item represents the amounts of work a worker can do in a day. Each worker is assigned to one job (bin). The benefit of a job is the minimum amount of work to do it, which is proportional to its size. This scenario relates to one dimensional bin covering.

The scenario of vector covering is equivalent to the case where a job has \(d\) different qualities. We can normalize and assume that each quality which exists in the job has size 1, each type of job has to have a subset of the qualities, the different types of jobs are again represented by bins, which are \(d\)-dimensional binary vectors. The work of each worker is also represented by a \(d\)-dimensional vector which corresponds to the parts of each quality that the worker can do in one day.

**Definition of competitive ratio**: We measure the performance of an on-line algorithm by the competitive ratio. Denote the value of the on-line algorithm by \(V_{on}\), and the value of an optimal off-line algorithm, that knows the sequence in advance, by \(V_{opt}\). The competitive ratio is the supremum \(r\) \((r \leq 1)\) for which \(V_{on} \geq r \cdot V_{opt} - \eta\) is satisfied for every sequence \((\eta\) is an additive constant which does not depend on the input). For randomized algorithms we replace \(V_{on}\) by \(E(V_{on})\) and in this case the competitive ratio is the supremum \(r\) for which
Let $B$ be derived with worst case ratios of $\frac{2}{3}$ and $\frac{3}{4}$, respectively. Both of these sophisticated algorithms are based on presorting the items, and consequently are off-line algorithms. The greedy algorithm, however, is an on-line algorithm. Csirik and Totik proved that in fact the greedy algorithm is a best possible on-line algorithm, since no on-line algorithm can have a worst case ratio that is strictly better than $1/2$. Csirik, Frenk, Galambos and Rinnooy Kan gave a probabilistic analysis of the one-dimensional bin covering and of the two-dimensional vector covering problem (for the case that all bins are of size $(1,1)$). Gaizer constructed an off-line approximation algorithm with worst case guarantee of $1/2$ for dimension $d = 2$. The article by Csirik and Frenk summarizes all results on vector covering problems that were derived till 1990. The multi-dimensional vector covering problem was studied by Alon, Azar, Csirik, Epstein, Sevastianov, Vestjens and Woeginger. The paper considers the case where the only allowed bin is the “all 1” vector. This paper considers both on-line and off-line vector covering. Finally Woeginger and Zhang studied variable sized one-dimensional vector covering. They define a ratio $r(B)$ which is the minimum between the minimum ratio of two consecutive bin sizes (ignoring all bins smaller than $1/2$) and the inverse of twice the size of the minimum sized bin among the bins which are larger or equal to $1/2$. Clearly, this number is smaller than $1/2$. They show this ratio $r(B)$ is optimal for a finite set of bins $B$. A survey on on-line packing and covering is given in [9].

**Our results:** We give some definitions in order to define the number $r(B,m)$. Let $B = \{b_1, \ldots, b_s\}$ be the set of allowed bin sizes, where $1 = b_1 > b_2 > \ldots > b_s$, and let $m$ be an integer $m > 1$ such that $1/m$ is an upper bound on item sizes. For each $1 \leq i \leq s$, let $b_{i,j} = b_i/j$ and let $B_i(m)$ be the set of fractions of $b_i$ between sizes $1/(2m)$ and $1/m$, that is $B_i(m) = \{b_{i,j} | 1 \leq j \leq 2m\} \cap [1/(2m), 1/m]$. We define a new set of bins by $C(m) = \cup_{1 \leq i \leq s} B_i(m)$. Enumerate the sizes of numbers in $C(m)$, $C(m) = \{c_1, \ldots, c_k\}$ where $1/m = c_1 > c_2 > \ldots > c_k = 1/(2m)$. For every element in $C(m)$, recall an original bin size that caused it to be inserted into $C(m)$. For $c_i$, let $b(c_i)$ be the smallest $b_j$ in a way that there exists an integer $y$ that satisfies $yc_i = b_j$. Now define $q(B,m) = \max\{c_i/c_{i+1} | 1 \leq i \leq k - 1\}$. Note that $1 + 1/m \geq q(B,m) > 1$. Finally define $r(B,m) = 1/q(B,m)$.

We show the following results for bin covering.
– For every finite set of bins $B$ and integer $m \geq 1$, we give a deterministic algorithm with competitive ratio $r(B, m)$.
– For every finite set of bins $B$ and integer $m \geq 1$, we give an upper bound of $r(B, m)$ on the competitive ratio of any algorithm.

These results show that $r(B, m)$ is the best possible competitive ratio for the set of bins $B$, and the upper bound $1/m$. This result reduces to the result of Woeginger and Zhang [11] for the case $m = 1$, and to the results of [3,8], if both $m = 1$ and $B = \{1\}$ hold. If only $B = \{1\}$ holds, then it follows from our result that the best competitive ratio is $m/(m + 1)$.

We show the following results for vector covering.
– An upper bound of $1/2^{d(1-o(1))}$ on the competitive ratio for the vector covering problem for arbitrary sets $B$.
– An algorithm with competitive ratio $1/2$ for the case where each bin in $B$ is a unit vector.
– An upper bound of $1/2$ on the competitive ratio for this case.
– For the case of unit prefix vectors, (i.e. bins are of the form $(1, \ldots, 1, 0, \ldots, 0)$, and $|B| = d$) we show that this model is harder than the case where $B$ consists of all unit vector bins (for large enough values of $d$) by giving an upper bound of $O(1/ \log d)$ on the competitive ratio of any algorithm.
– For the same case we show that the relation between models holds also for $d = 2$ by giving an upper bound of $4/9$ on the competitive ratio of any algorithm for $d = 2$.
– For the case where $B$ is the set of all possible binary vectors, we show that this model is easier than the basic model of one bin type which is all 1 (for large enough values of $d$) by giving an algorithm of ratio $\Omega(1/ \log d)$.
– For the same case we show that the relation between models holds also for $d = 2$ by giving an algorithm of ratio larger than $0.4$.

All upper bounds (negative results) in this paper hold both for deterministic and for randomized algorithms. In order to prove upper bounds we use an adaptation of Yao’s theorem [12] which states that an upper bound of a deterministic algorithm against a fixed distribution on the input is also an upper bound for randomized algorithms. For maximization problems the bound is given by $E(V_{on}/V_{opt})$ or by $E(V_{on})/E(V_{opt})$ [4,5]. All upper bounds are proved for algorithms that are allowed to have additive constants. In order to prove upper bounds for such algorithms, we show that all upper bounds hold for arbitrarily long sequences.

2 One-Dimensional Covering

In this section we show that for every set of one-dimensional bins $B$ and integer $m \geq 1$, the ratio $r(B, m)$ is the best possible competitive ratio.

Theorem 1. For every collection of bin sizes $B$, and parameter $m$, there exists an algorithm for bin covering with asymptotic worst case ratio $r(B, m)$. This result is best possible among all on-line randomized algorithms.
We start with the algorithm. Later we prove the upper bound which shows its optimality.

For a given set of bins $B$, and parameter $m$, we will use $q$ for $q(B, m)$, and $r$ for $r(B, m)$. Let $t = \lceil -\log_2 m(q-1) \rceil$. Define a partition of the interval of items $(0, 1/m)$ into the following subintervals:

- For $1 \leq j \leq k-1$ and $0 \leq l \leq t-1$, let $I_{j,l} = (c_{j+1}/2^l, c_j/2^l]$.
- Let $I_{k,t} = (0, c_k/2^{t-1}]$.

We assign each interval $I_{j,l}$ a corresponding bin size $d_{j,l}$. We define $d_{k,t}$ to be $b_1 = 1$. For each other interval $I_{j,l}$ we define $d_{j,k} = b(c_{j+1})$. The algorithm keeps one open bin of for each interval, this bin is used only for packing items in the interval. Each new item is classified and assigned to the corresponding bin. When a bin is covered, it is closed, and a new bin of the same size is opened and used for this interval.

**Lemma 1.** The competitive ratio of the above algorithm is at least $r$

**Proof.** We show that each covered bin is covered by at most $q$ times its size. For every first type interval $I_{j,l}$ we show that its corresponding closed bins are covered by at most $c_j/c_{j+1}$ times their size. Let $b(c_{j+1}) = yc_j$, then each such bin is covered by exactly $y2^l$ items of size at most $c_j/2^l$, and the total size of the items that cover it is at most $yc_j$. Thus the ratio between the size of items that cover the bin and the size of the bin is at most $c_j/c_{j+1}$ as required. Since $q \geq c_j/c_{j+1}$, the bin is covered by at most $q$ times its size.

For the interval $I_{k,t}$, since all items are bounded by $1/(2^t m)$, a covered bin is covered by items with total size at most $1+1/(2^t m)$. Since $2^t \geq 1/(m(q-1))$, $1+1/(2^t m) \leq q - 1 + 1 = q$, and each such bin is covered by at most $q$, which is $q$ times a size of a unit bin.

The total number of bins that are never closed is $O(kt)$, which is $O(mst)$ where $s = |B|$, this number depends solely on $B$ and $m$, and is an additive constant.

**Lemma 2.** The competitive ratio of any deterministic or randomized algorithm for bin covering is at most $r$.

**Proof.** We start with the deterministic case. Let $j$ be an integer such that $c_j/c_{j+1} = q$. Let $y$ be an integer such that $yc_j = b(c_j)$. Let $N$ be a large integer.

In order to define a list of items, we define a set of sizes $D$. Let $D = B \cup C'$ where $C' = \{xc' | x \in \{1, \ldots, 2m\}, c' \in C\} \cap [1/2m, 1]$. Enumerate the numbers in $D$, $D = \{d_1, \ldots, d_p\}$ where $d_1 > d_2 > \ldots > d_p$. Now define $\varepsilon > 0$ to be a number that satisfies $2m \cdot 2^N \varepsilon < b_s$ and for every $1 \leq j \leq p-1$, $2m \cdot 2^N \varepsilon < d_j - d_{j+1}$.

The list contains $2^N y$ items with size $\varepsilon$ (sand) followed by $2^{N-i} y$ items with size $c_j - 2^i \varepsilon$ for a fixed $i$, $0 \leq i \leq N$.

The optimal off-line assignment uses bins with size $b(c_j)$ to pack all items. Each bin contains $y$ items with size $c_j - 2^i \varepsilon$ and $2^i y$ items with size $\varepsilon$. This covers $2^{N-i}$ bins and $V_{opt} = 2^{N-i} b(c_j) = 2^{N-i} yc_j$. We change the packing of the online algorithm without reducing the value of the on-line algorithm in such a way
that the on-line algorithm will use only bins of weights $c_j$ and $c_{j+1}$. Consider a bin of the on-line algorithm of size $b$ which is covered, i.e. it contributes some amount to the on-line algorithm value. Since $y \leq 2m$, and $b \geq b_s > 2m2^N\varepsilon$, the bin contains at least one big item. Let $a > 0$ be the number of big items in the bin. Assume that $a \leq 2m$, otherwise remove items and let $a = 2m$, since items are larger than $1/2m$ and the bin is of size at most 1, then the bin is still covered.

Let $\mu$ be the total size of small items in the bin. If $\mu \geq 2^i a \varepsilon$ replace the bin by $a$ bins of size $c_j$ containing each one big item, and $\lfloor \mu/(a \varepsilon) \rfloor$ small items. Otherwise, partition the items in the same way, but use bins of weight $c_{j+1}$ instead.

We show that the new bins are covered and that the value of the on-line algorithm is not reduced, i.e. $b \leq ac_j$ in the first case and $b \leq ac_{j+1}$ in the second case. In the first case since $\mu/a \geq 2^i \varepsilon$, the total weight of items in a new bin is $(c_j - 2^i \varepsilon) + \lfloor \mu/(a \varepsilon) \rfloor \varepsilon \geq c_j$. Hence each new bin is covered. On the other hand $b \leq ac_j + y2^N \varepsilon$. According to the definition of $D$, $ac_j \in D$. Assume to the contrary that $b > ac_j$. Since $b \in D$ then $b \in D$, thus $b - ac_j > y2^N \varepsilon$ and hence the bin would not be covered. We conclude that $b \leq ac_j$, and the replacement of $b$ by $a$ bins of size $c_j$ does not change the on-line algorithm value and each new bin of size $c_j$ is covered by at least $c_j$. In the second case, if $\mu < 2^i a \varepsilon$, the weight of an item in a new bin is at least $c_j - 2^i \varepsilon > c_{j+1}$. Hence each new bin is covered. On the other hand $b < ac_j$. Assume to the contrary that $ac_{j+1} < b < ac_j$, then $b/a \in C$, contradicting the fact that $c_j$ is in the successor of $c_{j+1}$ in $C$. We conclude that new bins are covered, the on-line algorithm value was not reduced, and now the on-line algorithm uses only bins of sizes $c_j$ and $c_{j+1}$.

For $0 \leq i < N$, let $Z_i$ be the number of on-line algorithm bins with the amount of least $2^i \varepsilon$ weight of sand but less than $2^{i+1} \varepsilon$. Let $Z_N$ be the number of on-line algorithm bins with weight of sand at least $2^N$. We have $\sum_{i=0}^N Z_i 2^i \leq y2^N$. The on-line algorithm value would be $\sum_{i=0}^N Z_i 2^i \leq (y/2)((N + 2)c_{j+1} + 2c_j - 2c_{j+1}) = (yc_j/2^N)(N + 2)^{2^N - 1} = yc_j(N + 2)/2$.

For the optimal off-line algorithm we get $\sum_{i=1}^N 2^{i-1}/2^N \cdot 2^N - y2c_j = \sum_{i=1}^N 2^{i-1}/2^N \cdot 2N - y2c_j = (yc_j/2^N)(N + 2)^{2^N - 1} = yc_j(N + 2)/2$.

If the competitive ratio is $r$ and $q = 1/r$, then also for the convex sums $yc_j(N + 2)/2 \leq yq/(Nc_{j+1} + 2c_j)$ or $q \geq (N + 2)/((N(c_{j+1}/c_j) + 2)$. Hence $r = c_{j+1}/c_j + \delta_N$, where $\delta_N$ is arbitrarily close to 0, for large enough values of $N$.

For the randomized case we use the adaptation to Yao’s theorem, and the probabilities are the factors we multiplied by in the deterministic case. We get that $E(V_{on})/E(V_{opt}) \leq c_{j+1}/c_j$. 
3 Multi-dimensional Vector Covering

In this section we consider variable sized vector covering. In the classical vector covering problem, $B$ contains a single bin which is all 1. We study a simple extension of this model, and allow each bin to be a binary d-dimensional vector. We start by showing upper bounds on the competitive ratio for arbitrary sets $B$. The proof of the following Lemma is omitted.

**Lemma 3.** The competitive ratio of any deterministic or randomized on-line algorithm for binary variable sized vector covering is at most $1/2^d(1-o(1))$.

Due to this negative result, we do not study the general model, but restrict ourselves to specific interesting cases of $B$. We show some relations between the different models.

3.1 Covering a Single Type of Bin

We start with an easy model where $B$ consists of a single type of bin $b$. Assume that the number of non-zero components in $b$ is $i$, then we can reduce this problem to the case of vector covering bins of dimension $i$, where all components are 1 (the basic vector covering problem). This is true since both the on-line and off-line algorithms do not use the coordinates which are zero in the vector $b$. Thus the best competitive ratio for this case is $\Theta(1/i)$ [1].

3.2 Covering Unit Vector Bins

In this model we let $B$ consist of bins which have one non-zero component. We show that for any such set of bins, the best competitive ratio is 1/2. We consider this model since this is a case where all bins have the same bin weight.

The upper bound (negative result) follows from the bound of [8] by picking one $b \in B$, where the $i$th coordinate of $b$ is non-zero, and giving a sequence where all vectors are $d$-dimensional, with all coordinates zero, except for the $i$th coordinate which is the same as in [8]. We can also give an easier proof for the case where $|B| > 1$, we omit this proof due to space restrictions.

**Lemma 4.** If $B$ consists of unit vector bins only, and $|B| > 1$, then the competitive ratio of any deterministic or randomized on-line algorithm is at most 1/2.

We give an optimal greedy algorithm. We assume that $|B| = d$ and it contains all bins of unit bin weight. Otherwise we just cancel the coordinates that do not appear in any bin (are zero in all bins), since no algorithm can benefit from these coordinates, and treat the input as a lower dimension vector.

The algorithm has at most $d$ open bins at a time. Each arriving vector is classified to a type among $\{1, 2, \ldots, d\}$ according to its largest component.

Denote a bin whose $i$th coordinate is 1 by $b_i$. After an arriving item is classified to a type $j$, then if there is an open bin of size $b_j$, the item is assigned to
it, and if after the item is assigned to a bin, the bin is covered, we close the bin (and never use it again).

If no bin of size $b_j$ is open, open a new one and continue in the same way.

**Theorem 2.** The competitive ratio of the greedy algorithm is at least $\frac{1}{2}$.

**Proof.** For the optimal off-line algorithm, define by $S_{opt}$, the active area of items, that is, the area of items that was used to cover bins (i.e. the area that is not wasted, which is in the case of unit vectors, 1 for each covered bin). Clearly, $V_{opt} = S_{opt}$. Since all bins have only one non-zero coordinate, the active area of each item is in one coordinate. For an item $a$, let $a_{\max} = \max_{1 \leq i \leq d} a_i$. Clearly, $S_{opt} \leq \sum_{a \in A} a_{\max}$, where $A$ is the sequence of items.

On the other hand, consider the area $\sum_{a \in A} a_{\max}$, the on-line algorithm packed according to the maximum component.

Let $\alpha_i$ be the number of bins of type $b_i$ that the on-line algorithm managed to cover. For each such bin, the $i$th coordinate is covered by less than 2 (since at the time the last item was assigned there, the coordinate was less than 1, and the new item adds at most 1). There is at most one open bin which is never closed, and its $i$th coordinate is also less than 1. Hence $\sum_{a \in A_i} a_{\max} \leq 2\alpha_i + 1$, where $A_i$ are items that were classified to type $i$. Summing over all $1 \leq i \leq d$ we get: $\sum_{a \in A} a_{\max} \leq 2V_{on} + d$, which gives $V_{on} \geq \frac{1}{2}V_{opt} - \frac{d}{2}$.

### 3.3 Covering Unit Prefix Bins

We define a model which is similar to the previous one. Let $B_1$ be a set that contains exactly $d$ possible bin sizes. Bin $i$, which also has bin weight $i$, is a $d$-dimensional vector whose $i$ leftmost coordinates are 1, and all others are 0. We show that using this set of allowed bins instead of $d$ unit vector bins, makes the problem of on-line covering harder.

**Theorem 3.** The competitive ratio of any deterministic or randomized on-line algorithm for covering the increasing set of bins has competitive ratio of at most $O(1/\log d)$.

**Proof.** The upper bound sequence consists of $2k$ items with size $a = (1 - \epsilon, 0, 0, \ldots, 0)$ and then with probability $p_i$ (which we fix later), $2k$ items with size $(\epsilon, 1, \ldots, 1, 0, \ldots, 0)$ arrive ($d - i$ zeros), and for $2 \leq i \leq d$, with probability $q_i$, (which is also fixed later), $k$ items with size $(0, 1, \ldots, 1, 0, \ldots, 0)$ arrive ($d - i$ zeros). $q_1$ is the probability that no further items (except the first $2k$ items) arrive. Hence the probabilities should be fixed so that $\sum_{i=1}^{d} (p_i + q_i) = 1$.

Let $\alpha_i$ be the number of bins of bin weight $i$ that the on-line algorithm opens for one item of size $a$, and $\beta_i$ to be the number of bins of bin weight $i$ it opens, assigning two items with size $a$ to each.

Let us calculate the competitive ratio for each case: Consider the cases where items with $d - i$ rightmost zero coordinates arrive. With probability $p_i$, the on-line algorithm manages to cover all bins of weight of at most $i$ that it opened.

The optimal off-line assignment would use only bins of bin weight $i$, and would assign one item of each type to every bin. Thus $C_{opt} = 2ki$, and $C_{on} =$
\[ \sum_{j=1}^{i} (\alpha_j + \beta_j) \cdot j. \] On the other hand, with probability \( q_i \), the on-line algorithm would cover only bins with at least two items with size \( a \) (otherwise the leftmost component is not covered).

The optimal off-line assignment would use only bins of bin weight \( i \), and would assign two items with size \( a \), and one item of the other type to each bin. Thus \( C_{opt} = ki \) and \( C_{on} = \sum_{j=1}^{d} \beta_j \cdot j. \) According to the initial on-line assignment, we have \( \sum_{i=1}^{d} \alpha_i + 2\beta_i \leq 2k \). Let \( p_i = q_i = \frac{1}{2i(i+1)} \) for \( i < d \) and \( p_d = q_d = \frac{1}{2d}. \)

The expectation of the optimal off-line value is \( E(V_{opt}) = \sum_{i=1}^{d} (p_i 2ki + q_i ki) = 3k \left( \sum_{i=1}^{d-1} \frac{1}{2(i+1)} + \frac{1}{2} \right) = \frac{3}{2} k \left( \sum_{i=1}^{d-1} \frac{1}{i+1} + 1 \right) = \frac{3}{2} k \left( \sum_{i=1}^{d} \frac{1}{i} \right) \geq \frac{3}{2} k \ln d. \)

It is possible to show that the expectation of the on-line value satisfies \( E(V_{on}) \leq k \). The probabilities are valid since \( \sum_{i=1}^{d} (p_i + q_i) = \sum_{i=1}^{d-1} \frac{1}{i(i+1)} + \frac{1}{d} = \sum_{i=1}^{d-1} \left( \frac{1}{i} - \frac{1}{i+1} \right) + \frac{1}{d} = 1. \) Thus the competitive ratio is at most \( \frac{k}{2k \ln d} \leq \frac{2k}{3k \ln d} = O \left( \frac{1}{\log d} \right). \)

We show that this problem is significantly harder than unit vector bins, not only asymptotically but for \( d = 2 \) as well (in this case \( B = \{(1,1),(1,0)\} \)). The Lemma follows from a more careful examination of the general sequence for \( d = 2 \). The full proof is omitted.

**Lemma 5.** The competitive ratio of any deterministic or randomized algorithm for \( d = 2 \) and unit prefix bins is at most \( 4/9 \).

### 3.4 Covering All Possible Bins

In this model, \( B \) consists of \( 2^d - 1 \) bins, i.e. each binary \( d \)-dimensional bin may be used. We show that the basic covering model (the only allowed bin is the “all 1” bin) is harder than this model. We define the following algorithm.

Let \( k = \lceil \log_2(d^2) \rceil \). We define type vectors by vectors in the set \( A^d \) where \( A = \{0\} \cup \{\frac{1}{2^i} \mid 0 \leq i \leq k\} \). \( A^d \) consists of \((k+2)^d\) different type vectors. The algorithm keeps one open bin of a corresponding size for each vector type. To determine the bin vector for a certain type vector we do the following:

For each \( 0 \leq s \leq k \), let \( \alpha_s \) be the sum of sizes of all components of size \( \frac{1}{2^s} \) of the type vector, which is \( \frac{m}{2^s} \), where \( m \) is the number of \( \frac{1}{2^s} \) components. Let \( s_1 \) be an integer for which \( \alpha_{s_1} \) is maximized. The corresponding bin would have zeros in all coordinates which are less than \( \frac{1}{2^{s_1}} \) in the type vector, and 1 in all coordinates which are at least \( \frac{1}{2^{s_1}} \).

We keep one open corresponding bin for each possible type vector: On arrival of an item vector \( a \), \( a \) is identified with a type vector, then it is assigned to a bin that corresponds to this type. A covered bin is closed and replaced by a new bin of the same size.

For each arriving item vector \( a \), compute its type vector in the following way: Let \( i \) be the coordinate such that \( a_i = \max_{1 \leq j \leq d} a_j \). Let \( a' \) be the vector \( \frac{a}{a_i} \). Replace each coordinate \( a'_j \) by the maximum number \( \frac{1}{2^s} \), such that \( a'_j \geq \frac{1}{2^s} \). If \( s > k \), then replace \( a'_j \) by zero. The new vector \( a'' \) is the type vector of \( a \). Let
$a'''' = a'' \cdot a_i$. $a$ is assigned to the open bin which corresponds to the type vector $a''$.

**Theorem 4.** The above algorithm has competitive ratio $\Omega(\frac{1}{\log d})$.

The proof of the Theorem 4 is omitted. It is also possible to design an algorithm for $d = 2$, and prove the following theorem.

**Theorem 5.** There exists an algorithm for $d = 2$, for covering with all possible bins, which has competitive ratio strictly above $0.4$. The algorithm divides items into ranges, and packs each range separately.

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**References**

On Testing for Zero Polynomials by a Set of Points with Bounded Precision

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Abstract. We consider a general methodology proposed by Chen and Kao [4] for testing polynomial identities. We prove that the test cannot be completely derandomized by any specified set of rational approximations to algebraic numbers up to a polynomial number of bits. The proof is a direct application of Dirichlet’s box principle. We also give some number theoretic estimates for the likelihood of a multiplicatively independent sequence of integers which can be used in their algorithm.

1 Introduction

Randomization has been used very successfully in the field of algorithm design and complexity analysis. The first widely acclaimed use of randomization as a computational resource with provable polynomial time performance bounds can be traced to Solovay and Strassen [13] and to Rabin [11] following work by G. Miller [8]. These results were for testing primes. Another early work using randomness to achieve provable bounds in computation is the Schwartz-Zippel Theorem around 1978 [12] [14]:

Theorem 1 (Schwartz-Zippel). Let $F(x_1, \ldots, x_n) \in F[x_1, \ldots, x_n]$ be a multivariate polynomial of total degree $d$, where $F$ is any fixed field. Suppose $F$ is not identically zero. Then, for any fixed finite set $S \subseteq F$,

$$\Pr[F(r_1, \ldots, r_n) = 0] \leq \frac{d}{|S|},$$

where $r_1, \ldots, r_n$ are chosen independently and uniformly at random from $S$.

This elegant result is used to test for polynomial identities. Thus, if we wish to test for $f(x) = g(x)$, we simply pick a set of random inputs $r$ and check that $f(r) = g(r)$, with the understanding that if $f(x) = g(x)$, then certainly

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\( f(r) = g(r) \) for any \( r \), and, if \( f(x) \neq g(x) \), then \( f(r) \neq g(r) \) with high probability. The error probability is only on one side, and can be reduced by either choosing a larger set \( S \), or, as a more general principle, by running the probabilistic test many times. There are more advanced techniques for amplification (see [9,15]).

The Schwartz-Zippel method is a general technique for testing polynomial identities. Not only it can be used in straightforward algebraic settings, but also it can be used to give elegant probabilistic algorithms for combinatorial problems. A well known application is the probabilistic \( \text{NC}^2 \) algorithm for the perfect matching problem by Lovász [7].

A perfect matching in a graph \( G \) is a set \( M \) of edges, such that every vertex of \( G \) is incident to exactly one edge in \( M \). The perfect matching problem is to decide if a given graph has a perfect matching. Tutte gave a beautiful characterization of the existence of perfect matchings in a graph \( G \). Define the Tutte matrix \( T \) as follows. Suppose \( G \) has \( n \) vertices and \( m \) edges. Let \( \{x_{i,j}\} \) be a set of \( m \) indeterminants, one for each (undirected) edge \((i,j)\). For any edge \((i,j)\), with \( i < j \), put an \( x_{i,j} \) in the \((i,j)\) position and \(-x_{i,j}\) in the \((j,i)\) position, and 0 elsewhere.

**Theorem 2 (Tutte).** \( G \) has a perfect matching if and only if \( \det T \neq 0 \).

Then Lovász’s \( \text{RNC}^2 \) algorithm for the perfect matching problem is to apply the probabilistic test for polynomial identity and test for zero. More specifically, his algorithm assigns to each indeterminant a uniformly and independently chosen integer in the range \( \{1, \ldots, 2n\} \). Note that the total degree of \( \det T \) is (at most) \( n \), and thus the algorithm achieves success probability \( 1/2 \) with \( m \cdot \lceil \log_2(2n) \rceil \) random bits.

There are other randomized fast parallel algorithms for the perfect matching problem. For instance, Chari, Rohatgi and Srinivasan [3] gave an alternative algorithm that uses fewer random bits, but the processor count is a large polynomial in \( n \) (it is estimated that the algorithm by Chari et al needs to compute determinant and inverse of \( n \times n \) matrices of entries of size up to \( n^7 \); see [4].) We note that it has been a long standing open problem to achieve a deterministic \( \text{NC}^2 \) algorithm for the perfect matching problem.

Recently Chen and Kao [4] proposed a new general methodology for testing whether a polynomial with integer coefficients is identically zero. Their method is based on a theorem on the non-vanishing of a polynomial on algebraic numbers. Their motivation is to derandomize the \( \text{RNC}^2 \) algorithm for the perfect matching problem. The general idea of their approach is that, instead of evaluating the polynomial at random integer inputs, it can be evaluated at various algebraic numbers.

More specifically, let \( p_1 = 2, p_2, \ldots, p_n \) be the smallest \( n \) primes, and let \( f(x_1, \ldots, x_n) \) be a non-zero multilinear polynomial with integer coefficients. Then

\[
 f(\pm\sqrt{p_1}, \pm\sqrt{p_2}, \ldots, \pm\sqrt{p_n}) \neq 0.
\]

More generally, let \( f(x_1, \ldots, x_n) \) have degree \( d_i \) on \( x_i \). Let \( k_i = \lceil \log_2(d_i + 1) \rceil \), and \( s = \sum_{i=1}^n k_i \). Let \( q_{1,1}, \ldots, q_{1,k_1}, \ldots, q_{n,1}, \ldots, q_{n,k_n} \) be any \( s \) square-free positive
On Testing for Zero Polynomials by a Set of Points with Bounded Precision 475

integers which are pair-wise relatively prime, and let \( b_{1,1}, \ldots, b_{1,k_1}, \ldots, b_{n,1}, \ldots, b_{n,k_n} \) be any \( s \) bits from \( \{0, 1\} \). Then \( f \) is non-zero if and only if

\[
f \left( \sum_{j=1}^{k_1} (-1)^{b_{1,j}} \sqrt{q_{1,j}}, \sum_{j=1}^{k_2} (-1)^{b_{2,j}} \sqrt{q_{2,j}}, \ldots, \sum_{j=1}^{k_n} (-1)^{b_{n,j}} \sqrt{q_{n,j}} \right) \neq 0.
\]

This does give a deterministic test, albeit one which involves irrational numbers, and thus its exact computation is computationally infeasible.

To overcome this difficulty, Chen and Kao used a certain random combination of rational approximations to the irrational algebraic numbers. Indeed, in the above deterministic test there is no need for the \( s \) bits \( b_{i,j} \) at all; these were included for the purpose of using randomization. More specifically, they used a random choice of the \( s \) bits, with a truncation of the quadratic irrational numbers to a prespecified polynomial number of bits \( \ell \). They showed that to achieve good success probability, \( \ell = [\log_2 (c + cd)] + d[\log_2 n] + 2d[\log_2 s] + d + 1 \) suffices, where \( d \) is the total degree of \( f \) and \( c \) is its maximum absolute value of coefficients. In particular, they suggest using simply the truncations of \( (\pm \sqrt{p_1}, \pm \sqrt{p_2}, \ldots, \pm \sqrt{p_n}) \) for multilinear polynomials, where the signs \( \pm \) are randomly chosen.

With this method, they arrived at their randomized NC\(^2\) algorithm for the perfect matching problem. It turns out that the number of random bits used is only \( \sum_{i=1}^n [\log_2 \max\{1, n_i\}] \), where \( n_i \) is the number of neighbors \( v_j \) of the \( i \)-th vertex \( v_i \) in \( G \) such that \( j > i \). This allowed them to achieve an RNC\(^2\) algorithm that uses fewer random bits without doing more work than all previous randomized NC algorithms for the perfect matching problem. They also used their general methodology for several other problems.

In this paper we show that the use of prime square roots \( \sqrt{p_1}, \sqrt{p_2}, \ldots, \sqrt{p_n} \) is unnecessarily restrictive. Even if we use square-free positive integers which are pair-wise relatively prime, it is still unnecessarily restrictive. Instead we only need numbers that are multiplicatively independent modulo squares. Note that it is computationally easy to test for this property of multiplicative independence modulo squares \([1]\). Moreover, for polynomially many random integers, it is exponentially unlikely that they are pair-wise relatively prime. On the other hand, it is superpolynomially unlikely that they are multiplicatively dependent modulo squares. We prove these facts in Section 3.

While their algorithm reduces the number of random bits required, the possibility remains that perhaps one might be able to eliminate randomness altogether, by choosing carefully a specific set of signs in the rational combination of irrational numbers and a sufficiently good rational approximation to each of the irrational values, and then evaluating. As their methodology is a general one, the general question then is whether there are such universal choices. The authors of \([4]\) consider this the most intriguing problem left open from their approach \([5]\).

In Section 4 we show that this is in general impossible.
2 Algebraic Preliminaries

Let \( \mathbb{Q} \) be the field of rational numbers. Let \( \mathbb{Q}^* \) be the multiplicative group of \( \mathbb{Q} \), and \( \mathbb{Q}^{*2} \) be the subgroup of squares, i.e., \( \mathbb{Q}^{*2} \) is the image of \( x \mapsto x^2 \) from \( \mathbb{Q}^* \). By the unique factorization of integers, the quotient group \( R = \mathbb{Q}^*/\mathbb{Q}^{*2} \) is a direct sum of countably many copies of \( \mathbb{Z}_2 \) generated by \( p_0 = -1 \), and by \( p_1 = 2, \ldots, p_n, \ldots \), respectively, where \( p_n \) is the \( n \)-th prime, each of which has order 2 in \( R \).

Let \( \beta_1, \beta_2, \ldots, \beta_m \) be any \( m \) multiplicatively independent elements in \( R \), i.e., for any non-empty \( S \subseteq \{1, \ldots, m\} \), \( \prod_{i \in S} \beta_i \neq 1 \) in \( R \). Results of the following type are generally referred to as Kummer theory (see Lang [6] Chapter VIII Section 8). However we present the following elementary proof.

Lemma 1. Let \( Q_m = \mathbb{Q}(\sqrt{\beta_1}, \sqrt{\beta_2}, \ldots, \sqrt{\beta_m}) \). Then \( [Q_m : \mathbb{Q}] = 2^m \).

Proof. Write each \( \beta_j \) as a 0-1 vector of the exponents mod 2 over primes as follows. By unique factorization, \( \beta_j = \prod_{i \geq 0} p_i^{e_i} r_i^2 \), where \( r_2 \in \mathbb{Q}^{*2} \) and \( e_i \) mod 2 are uniquely defined. (Here \( p_0 = -1 \) and \( p_i, i \geq 1 \) are all the primes.) The uniqueness claim is just the elementary statement that for distinct primes \( p_i \), \( \pm \prod_{i \leq 1} p_i \neq 1 \) is not a square in \( \mathbb{Q}^* \). Then the 0-1 vector for \( \beta_j \) is \((e_0, e_1, \ldots)\). It is actually a finite vector, one entry for every prime appearing in \( \beta_j \); or one can think of it as an infinite 0-1 vector, with all but finitely many entries 0. This establishes an isomorphism between \( R \) and \( \oplus \mathbb{Z}_2 \). Now one can perform Gaussian elimination on the matrix over \( \mathbb{Z}_2 \) whose rows are these 0-1 vectors for \( \beta_1, \beta_2, \ldots, \beta_m \). (In matrix terms, the elementary row/column operations are: exchange of two rows or columns, adding mod 2 of one row to another.) Being multiplicatively independent is equivalent to the rank of this matrix being \( m \) over \( \mathbb{Z}_2 \). We can arrive at a normal form after Gaussian elimination \([I_m, B]\), where \( I_m \) is the \( m \times m \) identity matrix and \( B \) is some matrix over \( \mathbb{Z}_2 \) with \( m \) rows.

If \( \alpha_1, \alpha_2, \ldots, \alpha_m \) are the elements corresponding to the rows, then there are \( m \) distinct \( q_1, q_2, \ldots, q_m \in \{p_0, p_1, \ldots\} \), such that \( \alpha_j = q_j r_j \), where every \( r_j \) is a product of \( p_i \)'s other than \( q_1, q_2, \ldots, q_m \). In terms of field extension:

\[
Q_m = \mathbb{Q}(\sqrt{\alpha_1}, \sqrt{\alpha_2}, \ldots, \sqrt{\alpha_m}).
\]

Denote \( Q_j = \mathbb{Q}(\sqrt{\alpha_1}, \sqrt{\alpha_2}, \ldots, \sqrt{\alpha_j}) \).

Clearly \([Q_m : \mathbb{Q}] \leq 2^m \) as \([Q_j : Q_{j-1}] \leq [\mathbb{Q}(\sqrt{\alpha_j}) : \mathbb{Q}] = 2 \). There is an injective homomorphism from the Galois group \( G = \text{Gal}(Q_m/\mathbb{Q}) \) to the permutation group which maps \( \sqrt{\alpha_j} \) to \( \pm \sqrt{\alpha_j} \), for every \( j \). The full permutation group is the dual group of \( \mathbb{Z}_2^m \) which has order \( 2^m \). We want to show that this homomorphism is surjective, and thus \( G \) is isomorphic to the dual of \( \mathbb{Z}_2^m \) and in particular has order \( 2^m \) as well. Being the splitting field of \( \prod_{j=1}^m (X^2 - \alpha_j) \), \( Q_m \) is Galois over \( \mathbb{Q} \), and thus by the fundamental Galois correspondence \([Q_m : \mathbb{Q}] = |G| = 2^m \).

To show that \(|G| = 2^m \), we represent each \( \sigma \in G \) as a 0-1 vector \((\sigma_j)_{j=1}^m \) by the action on \( \{\sqrt{\alpha_j}\} \), \( \sigma(\sqrt{\alpha_j}) = (-1)^{\sigma_j} \sqrt{\alpha_j} \). Since each \( \alpha_j \) has a distinct \( q_j \), we have the following important property of \( G \): For any non-empty subset
of indices \( I \subseteq \{1, \ldots, m\} \), let \( \alpha_I = \prod_{j \in I} \alpha_j \), then there is an automorphism \( \tau \in \text{Gal}(\mathbb{Q}(\sqrt{\alpha_I})/\mathbb{Q}) \) which maps \( \sqrt{\alpha_I} \) to \( -\sqrt{\alpha_I} \). This \( \tau \) extends to some \( \sigma \) in \( G \). In terms of the 0-1 vectors \( (\sigma_j) \), for every non-empty subset \( I \) of \( \{1, \ldots, m\} \), there is some \( (\sigma_j) \), such that \( \sum_{j \in I} \sigma_j = 1 \mod 2 \).

We claim that the 0-1 vectors obtained from \( G \) include \( m \) unit vectors where the \( i \)-th unit vector has its \( i \)-th entry 1 and 0 elsewhere. Inductively, we assume there are \( v^{(i)} \), \( 1 \leq i \leq j \), whose restriction to the first \( j \) coordinates are the first \( j \) unit vectors. The base case is established by taking \( I = \{1\} \). For \( j + 1 \), take \( I = \{i \mid 1 \leq i \leq j; \ v^{(i)}_{j+1} = 1\} \cup \{j + 1\} \subseteq \{1, 2, \ldots, j + 1\} \). Then there is a \( u \) from \( G \), such that \( \sum_{k \in I} u_k = 1 \mod 2 \). Define

\[
v^{(j+1)} = u + \sum_{i:1 \leq i \leq j, u_i=1} v^{(i)},
\]

then \( v^{(j+1)} \) satisfies our requirement.

This is fairly clear. Formally, starting with \( k \leq j \),

\[
v^{(j+1)}_k = u_k + \sum_{i:1 \leq i \leq j, u_i=1} v^{(i)}_k = u_k + \sum_i 1_{[u_i \equiv 1, i=0]} = 0 \mod 2.
\]

For coordinate \( j + 1 \),

\[
v^{(j+1)}_{j+1} = \sum_{k \in I} v^{(j+1)}_k \quad \text{(adding zero terms)}
= \sum_{k:1 \leq k \leq j, k \in I, u_k=1} (u_k + v^{(k)}_{j+1}) + u_{j+1}
= \sum_{k \in I} u_k
= 1 \mod 2.
\]

To complete the induction, we replace \( v^{(i)} \) by \( v^{(i)} + v^{(j+1)} \) for \( i \in I, i \leq j \).

We actually established more, namely we computed the Galois group \( G = \text{Gal}(\mathbb{Q}_m/\mathbb{Q}) \). Moreover, reverting back to the multiplicatively independent set \( \{\beta_1, \ldots, \beta_n\} \), for any \( j \), the proof gives a permutation \( \sigma \in G \) which flips \( \sqrt{\beta_j} \) to \( -\sqrt{\beta_j} \) but fixes all others. These generate \( G \) as a free abelian group of exponent 2. Hence we have

**Corollary 1.** For any sequence \( a_1, a_2, \ldots, a_m \in \mathbb{Q}^* \), \( \mathbb{Q}(\sum_{i=1}^m a_i \sqrt{\beta_i}) = \mathbb{Q}_m \). In particular this is true for any sign sequences \( a_i = \pm 1 \).

**Proof.** If \( \mathbb{Q}(\sum_{i=1}^m a_i \sqrt{\beta_i}) \) were a proper subfield of \( \mathbb{Q}_m \), there would be a non-trivial automorphism of \( G \) which fixes \( \sum_{i=1}^m a_i \sqrt{\beta_i} \). But a non-trivial automorphism must flip some \( \sqrt{\beta_i} \), leading to a non-trivial linear relation \( \sum_S 2a_i \sqrt{\beta_i} = 0 \), where \( S \neq \emptyset \). Knowing the structure of \( G \), this is obviously impossible as we can apply some \( \sigma \) which flips one \( \sqrt{\beta_{i_0}} \) where \( i_0 \in S \) and fixes all others, and get \( \sqrt{\beta_{i_0}} = 0 \).
Let $f(x_1,\ldots,x_n) \in \mathbb{Q}[x_1,\ldots,x_n]$ have degree $d_i$ on $x_i$. Let $k_i = \lfloor \log_2(d_i + 1) \rfloor$, and $s = \sum_{i=1}^n k_i$. Let $q_{1,1},\ldots,q_{1,k_1},\ldots,q_{n,1},\ldots,q_{n,k_n}$ be any $s$ multiplicatively independent elements of $R=\mathbb{Q}^*/\mathbb{Q}_{\mathbb{Q}}^2$, and $a_{1,1},\ldots,a_{1,k_1},\ldots,a_{n,1},\ldots,a_{n,k_n}$ be any $s$ elements of $\mathbb{Q}^*$. Denote by $q_i = \sum_{j=1}^{k_i} a_{i,j} \sqrt{q_{i,j}}$, then

**Corollary 2.** We have $\mathbb{Q}(q_1,\ldots,q_n) = \mathbb{Q}(\sqrt{q_{1,1}},\ldots,\sqrt{q_{1,k_1}},\ldots,\sqrt{q_{n,1}},\ldots,\sqrt{q_{n,k_n}})$, and $f$ is non-zero if and only if $f(q_1,\ldots,q_n) \neq 0$.

**Proof.** The first equality follows from Corollary 1 by noting the following sandwiched containment

$$\mathbb{Q}(\sqrt{q_{1,1}},\ldots,\sqrt{q_{1,k_1}},\ldots,\sqrt{q_{n,1}},\ldots,\sqrt{q_{n,k_n}}) = \mathbb{Q}(\sum_{i=1}^n q_i) \subseteq \mathbb{Q}(q_1,\ldots,q_n) \subseteq \mathbb{Q}(\sqrt{q_{1,1}},\ldots,\sqrt{q_{1,k_1}},\ldots,\sqrt{q_{n,1}},\ldots,\sqrt{q_{n,k_n}}).$$

Hence for all $j \geq 1$

$$[\mathbb{Q}(q_1,\ldots,q_j) : \mathbb{Q}(q_1,\ldots,q_{j-1})] = 2^{k_j}.$$

The second claim follows from this by a simple induction. \(\square\)

The algorithm by Chen and Kao [4] uses the special case of Corollary 2 on pair-wise relatively prime integers to arrive at the test for polynomial identities. With Corollary 2 we have generalized their algorithm to be used with any multiplicatively independent rationals or integers. In the next section we estimate the probability that polynomially many random integers are multiplicatively dependent modulo squares. We show that this probability is superpolynomially small. Thus most such sequences of integers are multiplicatively independent. On the other hand, the probability that polynomially many random integers are pair-wise relatively prime is exponentially small.

### 3 The Probability of Multiplicative Dependence

In this section we consider various probabilities related to multiplicative dependence.

We will first show that the probability that a multiplicative dependence modulo squares among polynomially many random integers is extremely small. For this result we need the function $L(N) := \exp(\log N \log \log N)$. Pomerance [10] showed that if we choose random positive integers $q_1,\ldots,q_n \leq N$, then (from our point of view)

$$\Pr[\deg \mathbb{Q}(\sqrt{q_1},\ldots,\sqrt{q_n}) < 2^n] \rightarrow \begin{cases} 0, & \text{if } n \leq L(N)^{\frac{3}{2}-\epsilon}, \\ 1, & \text{if } n \geq L(N)^{\frac{3}{2}+\epsilon}. \end{cases}$$

His bound for the probability in the first case is $O(1/\sqrt{\log N})$, which is not sharp enough. We are, however, interested in much smaller $n$ relative to $N$, and can simplify his argument and still get a useful bound.
Theorem 3. Fix $d > 0$. Let $q_1, q_2, \ldots, q_n$ be chosen independently from the uniform distribution on $\{1, \ldots, N\}$, with $n \leq (\log N)^d$. The probability that there is a nonempty $S \subseteq \{1, \ldots, n\}$ and an integer $r$ such that $\prod_{i \in S} q_i = r^2$ is at most $L(N)^{-1/\sqrt{2} + o(1)}$.

Proof. We will call a number $B$-smooth if all its prime factors are less than or equal to $B$. If a nontrivial product from $q_1, \ldots, q_n$ is a square, then at least one of the following two events must occur. Either for some prime $p > B$, we have $p^2 | q_i$ for some $i$ or $p | q_i, q_j$ for some $i < j$ (call this $C$), or one of $q_1, \ldots, q_n$ is $B$-smooth (call this $S$). We have $\Pr[C] \leq \sum_{p > B} (n/p^2 + (n/2)/p^2) \leq (n+1)/B$. Also, $\Pr[S] \leq n\psi(N,B)/N$, where $\psi(N,B)$ counts the $B$-smooth numbers $\leq N$. From [2] we know that when

$$u := \frac{\log N}{\log B} \leq \frac{\log N}{\log \log N},$$

we have $\psi(N,B)/N \leq e^{-u \log u(1+o(1))}$, uniformly as $u \to \infty$. The choice of $B$ involves a tradeoff that is familiar from the analysis of factorization algorithms: we would like $B$ large so as to make $C$ unlikely, but we would also like it small so as to prevent $S$. A good choice is

$$B = \exp \left( \sqrt{\frac{\log N \log \log N}{2}} \right), \quad u = \sqrt{\frac{2 \log N}{\log \log N}},$$

which makes the probabilities roughly equal. Then it is straightforward to prove that $\Pr[C]$ and $\Pr[S] \leq L(N)^{-1/\sqrt{2} + o(1)}$, from which the result follows. \hfill \square

Note that the bound $L(N)^{-1/\sqrt{2} + o(1)}$ is asymptotically much less than inverse exponential polylog in input size $\log N$.

Corollary 3. For any $c > 0$, under the hypotheses of the last theorem, we have

$$\Pr[\deg Q(\sqrt{q_1}, \ldots, \sqrt{q_n}) < 2^n] = o(1/\exp((\log \log N)^c))$$

as $N \to \infty$.

The probability that $q_1, \ldots, q_n$ are pairwise relatively prime is also of interest. It is easy to see that the probability of this is exponentially small, for example by considering adjacent pairs. Below we give a bound with explicit dependence on the parameters.

Theorem 4. Let $q_1, \ldots, q_n$ be chosen independently from the uniform distribution on $\{1, \ldots, N\}$. The probability that these numbers are pairwise relatively prime is at most $(1 + n)2^{-n} + n/N$.

Proof. For the $q_i$ to be pairwise relatively prime, it is necessary and sufficient that this condition hold locally for each $p$. To get our estimate we use only $p = 2$. First assume that $N$ is even. The residues of the $q_i \mod 2$ are a random sequence of bits, of which at most one can be 0. Using the binomial distribution,
the probability of this event is $2^{-n} + n 2^{-n} = (1 + n)2^{-n}$. If $N$ is odd, either all $q_i$ are $\leq N - 1$, for which the previous case applies, or some $q_i$ equals $N$. The chance of this last event is at most $n/N$. 

One can get better results by considering more primes.

**Theorem 5.** For $n$ fixed and $N \to \infty$, the probability of the last theorem has the limit

$$P(n) = \prod_p \left\{ \left(1 + \frac{n}{p-1}\right) \left(1 - \frac{1}{p}\right)^n \right\}.$$ 

**Proof.** Let $E_N$ be the event that $q_1, \ldots, q_n$ are pairwise relatively prime. Let $G_p$ ("good at $p$") be the event that no pair has a gcd divisible by $p$, and $B_p$ ("bad at $p$") its complement. Then

$$\Pr \left[ \bigcap_{p \leq x} G_p \right] - \sum_{p > x} \Pr[B_p] \leq \Pr[E_N] \leq \Pr \left[ \bigcap_{p \leq x} G_p \right].$$

We have $\sum_{p > x} \Pr[B_p] \leq \binom{n}{2} x^{-1}$, so

$$P_x(n) - \frac{\binom{n}{2}}{x} \leq \liminf_{N \to \infty} \Pr[E_N] \leq \limsup_{N \to \infty} \Pr[E_N] \leq P_x(n),$$

where

$$P_x(n) = \prod_{p \leq x} \left(1 + \frac{n}{p-1}\right) \left(1 - \frac{1}{p}\right)^n.$$ 

By the binomial theorem, each factor of $P_x$ is $1 + O(p^{-2})$, so $\lim_{x \to \infty} P_x = P(n)$, which implies the theorem. 

### 4 The Lower Bound

In this section we prove that for any choice of $s$ multiplicatively independent $q_{i,j}$’s of polynomially many bits, where $1 \leq i \leq n$, $1 \leq j \leq k_i = \lceil \log_2(d_i + 1) \rceil$, $s = \sum_{i=1}^n k_i$, and any prespecified choice of signs in the sums $q_i = \sum_{j=1}^{k_i} (-1)^{b_{i,j}} \sqrt{q_{i,j}}$, and any polynomial bit precision $L$, there are non-zero polynomials $f(x_1, \ldots, x_n)$, where the degree of $x_i$ is at most $d_i$ and the coefficients are integral and all bounded in $n$, such that while $f(q_1, \ldots, q_n) \neq 0$, for all $\ell \leq L$, $f([q_1]_{\ell}, \ldots, [q_n]_{\ell}) = 0$, where $[q_i]_{\ell}$ denotes the $\ell$-bit truncation of $q_i$. This provides a negative answer to the open question from Chen and Kao [4]. We can even take $f$ multilinear and all coefficients from $\{-1, 0, 1\}$.

In fact, a much stronger result can be shown where $L$, the bit precision, can be substantially increased from polynomial in $n$ to almost $2^n$, yet still some non-zero multilinear polynomial $f$ with coefficients from $\{-1, 0, 1\}$ can vanish on all truncations up to $L$ bits. Such a polynomial can even simultaneously
vanish on all $L$-bit truncations on a large (almost $2^n$) set of arbitrarily chosen initial irrational points. The proof is a direct adaptation of Dirichlet’s method showing the existence of simultaneous rational approximations with bounded denominators.

**Theorem 6.** Given any $n$ and any degree sequence $d_1, \ldots, d_n$, let $D = \prod_{i=1}^n \bigl( d_i + 1 \bigr)$. Let $C, L, Q$ be arbitrary positive integers, such that $(2C+1)^D > (2^{L+1}CDQ)^L$. Let $q_{1,1}, \ldots, q_{1,k_1}, \ldots, q_{n,1}, \ldots, q_{n,k_n}$ be any integers multiplicatively independent modulo squares, where $k_i = \lceil \log_2(d_i + 1) \rceil$. Let $q_i = \sum_{j=1}^{k_i} a_{i,j} q_{i,j}$, where $a_{i,j}$ are any non-zero integers. Assume $\prod_i |q_i|^{d_i} \leq Q$. Then, there are non-zero polynomials $f(x_1, \ldots, x_n)$ with integral coefficients bounded by $C$, and degree in $x_i$ bounded by $d_i$, such that $f([q_1]_{\ell}, \ldots, [q_n]_{\ell}) \neq 0$, for all $1 \leq \ell \leq L$, where $[q_i]_{\ell}$ denotes the $\ell$-bit truncation of $q_i$.

**Corollary 4.** If $2C = 2^c$, $Q = 2^d$, and $2^n \geq L(L + c + q + n)/c$, then there are non-zero multilinear polynomials satisfying the above condition.

**Corollary 5.** If $2^n \geq tL(L + q + n + 1)$, then there are non-zero multilinear polynomials with coefficients from $\{-1, 0, 1\}$ that vanish on all $L$-bit truncations on a set of $t$ arbitrarily chosen initial irrational points $q^{(1)}, q^{(2)}, \ldots, q^{(t)}$, where each $q^{(i)} = (q^{(i)}_1, q^{(i)}_2, \ldots, q^{(i)}_n)$ has the form given in Theorem 6.

**Proof.** For any $\xi = (\xi_{j_1}, \ldots, \xi_{j_n})$, where $\xi_{j_1}, \ldots, \xi_{j_n}$ is a sequence of integers, with $0 \leq j_i \leq d_i, 1 \leq i \leq n$ and $|\xi_{j_1}, \ldots, \xi_{j_n}| \leq C$, define a polynomial $f_\xi$ by

$$f_\xi(x_1, \ldots, x_n) = \sum_{0 \leq j_1 \leq d_1, \ldots, 1 \leq i \leq n} \xi_{j_1} \cdots x_n^{j_n}.$$

The evaluations of $f$ at all $L$-bit truncations of $q_i$ defines a (linear) map from $[-C, C]^D$ to $(-CDQ, CDQ)^L$, by $\xi \mapsto (f_\xi([q_1]_{\ell}, \ldots, [q_n]_{\ell}), 1 \leq \ell \leq L)$. Note that there are at most $D$ non-zero terms in $f_\xi$ and each term is strictly less than $CQ$, being strict since $q_i$ is irrational.

Now divide the range into $2^{L+1}CDQ$ equal parts along every dimension. More precisely, for each dimension $1 \leq i \leq L$, divide the interval range $(-CDQ, CDQ)$ into $2^{L+1}CDQ$ equal length subintervals, i.e., for $1 \leq j_i < 2^{L+1}CDQ$ the $j_i$th subinterval is $(-CDQ + (j_i - 1)/2^L, -CDQ + j_i/2^L]$ and for the last one $j_i = 2^{L+1}CDQ$, the subinterval is $(CDQ - 1/2^L, CDQ)$. In this way, we obtain $2^{L+1}CDQ$ cells each congruent to $(0, 1/2^n]^L$ (except on each dimension the last one has open boundaries on both sides).

Since $(2C + 1)^D > (2^{L+1}CDQ)^L$, some two images fall in the same cell, i.e., there are two distinct tuples $\xi \neq \xi'$ such that

$$|f_\xi([q_1]_{\ell}, \ldots, [q_n]_{\ell}) - f_{\xi'}([q_1]_{\ell}, \ldots, [q_n]_{\ell})| < \frac{1}{2^L},$$

for all $1 \leq \ell \leq L$. 

However, each quantity is a rational number with at most $L$ bits after the decimal point. In other words, both $2^L f_{\xi}([q_1]|_\ell, \ldots, [q_n]|_\ell)$ and $2^L f_{\xi'}([q_1]|_\ell, \ldots, [q_n]|_\ell)$ are integers. It follows that $f_{\xi}([q_1]|_\ell, \ldots, [q_n]|_\ell) - f_{\xi'}([q_1]|_\ell, \ldots, [q_n]|_\ell) = 0$, for all $1 \leq \ell \leq L$. Now take $z_{j_1, \ldots, j_n} = \xi_{j_1, \ldots, j_n} - \xi'_{j_1, \ldots, j_n}$, then the non-zero polynomial $f_z$ has $f_z([q_1]|_\ell, \ldots, [q_n]|_\ell) = 0$, for all $1 \leq \ell \leq L$. \hfill \Box

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A Randomized Algorithm
for Gossiping in Radio Networks

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Abstract. We present an $O(n \log^4 n)$-time randomized algorithm for gossiping in radio networks with unknown topology. This is the first algorithm for gossiping in this model whose running time is only a polylogarithmic factor away from the optimum. The fastest previously known (deterministic) algorithm for this problem works in time $O(n^{3/2} \log^2 n)$.

1 Introduction

The two classical problems of disseminating information in computer networks are broadcasting and gossiping. In broadcasting, we want to distribute a given message from a distinguished source node to all other nodes in the network. In gossiping, each node $v$ in the network initially contains a message $m_v$, and we wish to distribute each message $m_v$ to all nodes in the network. In both problems, we would like to minimize the time needed to complete the task.

In radio networks, a message transmitted by a processor is sent to all processors within its range. The range relation is represented by a graph of nodes with directed edges between them. All processors work synchronously, and if a processor $u$ transmits a message $m$ at time step $t$, the message reaches each neighbor $v$ of $u$ at the same time step. Node $v$ will successfully receive $m$ only if $u$ is the only processor, among those whose range contains $v$, that transmits at time $t$. Since the communication links are uni-directional, there is no feedback mechanism in the network (see, for example, \cite{20}), and thus, in general, a node does not know for certain whether its transmissions were successful. Further, we assume that collisions cannot be resolved nor detected, that is, if messages from two or more processors reach $v$ at time $t$, $v$ does not get any message and it does not know that the collision occurred. This is motivated by situations when message collisions are difficult to distinguish from background noise, in which protocols that do not depend on the accuracy of the collision detection mechanism (see \cite{13,12}) will be more reliable.

We focus on gossiping algorithms that do not use any information about network topology. Such topology-independent algorithms are useful in networks...
with mobile users or unstable topologies, since then one does not need to change or reconfigure the protocol after topology changes. As long as the network is strongly connected and no changes occur during the actual execution of the algorithm, the task of gossiping will complete successfully. The strong connectivity assumption is necessary for gossiping to be meaningful.

**Past work.** Most of the previous work on radio networks focussed on broadcasting. If the topology of the network is known to all processors, Gaber and Mansour [14] showed that broadcasting can be achieved in time $O(D + \log^5 n)$, where $D$ is the network diameter. Diks et al [10] gave efficient broadcasting algorithms for special types of known networks. It is also known that computing an optimal broadcast schedule for a given network is NP-hard, even for points in the plane, where the graph is induced by node ranges, see [8,26].

For networks with unknown topology, Bar-Yehuda et al [3] gave a randomized algorithm that achieves broadcast in expected time $O(D \log n + \log^2 n)$. This is very close to the lower bound of $\Omega(D \log(n/D))$, by Kushilevitz and Mansour [18], and it matches this lower bound for a wide range of depth values, for example when $D = \Theta(n^{1-\epsilon})$, for any $\epsilon > 0$. Further, if $D$ is a constant, it also matches the lower bound of $\Omega(\log^2 n)$ for constant diameter networks, obtained by Alon et al [1].

In the deterministic case, Bar-Yehuda et al [3] gave an $\Omega(n)$ lower bound for constant diameter networks. For general networks, the best currently known lower bound of $\Omega(n \log n)$ was obtained by Bruschi and del Pinto [5], and, independently by Chlebus et al [6]. In [6], the authors also present a broadcast algorithm with time complexity $O(n^{11/6})$ – the first sub-quadratic upper bound. This upper bound was later improved to $O(n^{5/3} \log^3 n)$ by De Marco and Pelc [11]. Chlebus et al [7] developed several broadcasting algorithms, including one with time complexity $O(n^{3/2})$. In an independent work, using a probabilistic construction, Peleg [23] gave an $O(n^{3/2} \sqrt{\log n})$ upper bound. Recently, Chrobak, Gąsieniec and Rytter [9] presented a deterministic algorithm for broadcasting with time complexity $O(n \log^2 n)$, thus nearly matching the lower bound of $\Omega(n \log n)$ from [5,6]. As in [23], this broadcasting algorithm was constructed using a probabilistic argument.

The problem of gossiping has been intensely studied in various network models (see, for example, [15]). For radio networks, Ravishankar and Singh [24,25] studied gossiping algorithms for some restricted topologies, including paths and rings, under probabilistic assumptions on the spatial distribution of nodes. In our previous work, [9], we developed a deterministic algorithm for gossiping with time complexity $O(n^{3/2} \log^2 n)$, which, to our knowledge, is the only sub-quadratic algorithm for gossiping in radio networks with unknown topology.

**Our results.** In this paper we give a randomized $O(n \log^4 n)$-time algorithm for gossiping in radio networks with unknown topology. Our basic algorithm is Monte Carlo and it has the following performance characteristics: for any $0 < \epsilon < 1$, in time $O(n \log^3 n \log(n/\epsilon))$ it achieves broadcast with probability at least $1 - \epsilon$. This easily yields a Las Vegas algorithm with expected running time $O(n \log^4 n)$. 

2 Preliminaries

Radio networks. A radio network (see [3, 7]) is defined as an n-node directed graph whose nodes are assigned unique identifiers from the set \{1, 2, \ldots, n\}. Throughout the paper, for gossiping to be meaningful, we assume that the network is strongly connected. If there is an edge from \(u\) to \(v\), then we say that \(v\) is an out-neighbor of \(u\) and \(u\) is an in-neighbor of \(v\).

Initially, each node \(v\) contains a message \(m_v\) and has no other information. The time is divided into discrete time steps. All nodes start simultaneously, have access to a common clock, and work synchronously. (As noted by Peleg [23], the assumption about a common clock is not necessary.) At any time step, a node can be in one of two states: the receiving state or the transmitting state. A gossiping algorithm is a protocol that for each identifier \(id\) and for each time step \(t\), given all past messages received by \(id\), specifies the state of \(id\) at time \(t\). If \(id\) transmits at time \(t\), the protocol specifies the message. A message \(m\) transmitted at time \(t\) from a node \(u\) is sent instantly to all its out-neighbors. However, an out-neighbor \(v\) of \(u\) receives \(m\) at time step \(t\) only if \(v\) is in the receiving state and if no collision occurred, that is, if the other in-neighbors of \(v\) do not transmit at time \(t\) at all. Further, collisions cannot be distinguished from background noise. If \(v\) does not receive any message at time \(t\), it knows that either none of its in-neighbors transmitted at time \(t\), or that at least two did, but it does not know which of these two events occurred.

The running time of a gossiping algorithm is the smallest \(t\) such that for any strongly connected network topology, and for any assignment of identifiers to the nodes, each node receives all messages \(m_v\) no later than at step \(t\).

Simplifying assumptions. For clarity of presentation, we will present our algorithms as if the nodes knew \(n\), the size of the network. This assumption can be eliminated by a standard doubling technique (see [6, 7]) that works as follows: We organize the computation into phases, and we modify a given algorithm so that in phase \(i\) only nodes with labels at most \(2^i\) participate in the algorithm. This does not change the asymptotic running time.

Further, we will also assume throughout the paper that \(n\) is a power of 2. For other \(n\), the processors can execute the algorithm for the nearest power of 2 larger than \(n\), without changing the asymptotic running time.

Notation. By \(V\) we denote the set of nodes, and individual nodes are denoted by letters \(u, v, \ldots\). Messages are denoted by letter \(m\), possibly with indices. The message originating from a node \(v\) is denoted by \(m_v\). The whole set of initial messages is \(M = \{m_v : v \in V\}\). During the computation, each node \(v\) will store a set of messages \(M_v\) that have been received by \(v\) so far. Initially, \(M_v = \{m_v\}\). Without loss of generality, whenever a node is in the transmit mode, we can assume that it transmits the whole \(M_v\). This is achieved by a procedure denoted \text{transmit}(M_v). Procedure \text{receive}() puts \(v\) in the receive mode and it returns the received message, or \text{null} if no message has been received.
3 Limited Broadcast

One component of our algorithm is a procedure for limited broadcasting. Given an integer \( k \) and a node \( v \), the goal of limited broadcasting is to send the message \( M_v \) to at least \( k \) nodes in the network. We refer to \( v \) as the source node or the node that initiates the broadcast, and to \( M_v \) as the source message.

In [9], the broadcasting algorithm is defined by a sequence \( S = S_0S_1 \ldots \) of subsets of \( \{1, 2, \ldots, n\} \). At time \( t \), any node \( v \) that has already received the source message checks whether \( v \in S_t \). If so, \( v \) transmits the message, otherwise \( v \) is quiet. We modify the algorithm from [9], so that it performs the broadcasting procedure for only \( O(k \log^2 n) \) steps. Below, we appropriately refine the correctness proof, since the proof from [9] is not sufficient for our purpose.

The pseudo-code for the algorithm executed by each node is given below. Each node \( v \) has a boolean flag \( \text{active}_v \) that indicates whether \( v \) is active, that is, whether \( v \) has received a source message. Each iteration of the for-loop lasts one time step. The value of constant \( \gamma \) will be determined later.

Procedure \( \text{LtdBroadcast}_v(k) \).

\[
\text{for } \tau = 0, 1, \ldots, \gamma k \log^2 n - 1 \text{ do}
\begin{align*}
\text{if } \text{active}_v \text{ and } v \in S_\tau & \text{ then transmit}(M_v) \\
\text{else} & \\
\quad m & \leftarrow \text{receive}(); \\
\quad \text{if } m \neq \text{null} & \text{ then} \\
\quad\quad M_v & \leftarrow M_v \cup \{m\}; \\
\quad\quad \text{active}_v & \leftarrow \text{true}
\end{align*}
\]

Lemma 1. Assume that initially exactly one node \( u \) is active and that all nodes \( v \) begin executing \( \text{LtdBroadcast}_v(k) \) simultaneously. Then, for some constant \( \gamma \) independent of \( k \), after the computation is complete at least \( k \) nodes will receive the message from \( u \).

Proof. Without loss of generality, assume \( n \geq 2 \). In [9], it was shown that there is a constant \( \alpha \) such that for each \( j = 0, \ldots, \log n \) there is a family \( \tilde{S}_j = (S_{j,0}, S_{j,1}, \ldots, S_{j,a_j - 1}) \) of \( a_j = \alpha 2^j \log n \) sets with the following property:

\((*)\) For any two disjoint sets \( X, Y \) with \( 2^{j-1} \leq |X| \leq 2^j \) and \( |Y| \leq 2^j \) there exists a set \( S_{j,i} \) in \( \tilde{S}_j \) such that \( |X \cap S_{j,i}| = 1 \) and \( Y \cap S_{j,i} = \emptyset \).

Let \( \gamma = 12\alpha + 2 \). The sequence \( \tilde{S} \) consists of stages, with each stage, except possibly the last one, having \( \log n + 1 \) steps. Note that the number of stages is \( \lceil \gamma k \log^2 n/(\log n + 1) \rceil \geq 6a_k \log n + 1 \). The transmission set at the \( j \)th step of stage \( s \), that is \( S_\tau \) for \( \tau = s(\log n + 1) + j \), is \( S_{j,s \mod a_j} \).

Among the active nodes we distinguish two types of nodes: frontier nodes, which still have inactive out-neighbors, and inner nodes, which don’t. If there is a time step \( t < \gamma k \log^2 n \) when there are \( k \) or more frontier nodes, then the lemma holds trivially. So from now on we assume that at each step the number of frontier nodes is less than \( k \).
We define a sequence of stages \( s_0 = 0, s_1, \ldots, s_{l+1} = 6\alpha k \log n \), where \( 1 \leq s_{c+1} - s_c \leq 2\alpha k \log n \) for all \( c \). Denote by \( i_c \) and \( f_c \) the number of inner and frontier nodes when stage \( s_c \) is about to start. We will choose \( s_1, \ldots, s_l \) so that the following invariant holds for each \( c \leq l \):

\[
2i_c + f_c \geq \frac{s_c}{2\alpha \log n}.
\]  

(1)

Given (1), we can prove the lemma as follows. The number of nodes that have received the message when stage \( s_{l+1} \) ends is at least

\[
i_l + f_l \geq \frac{1}{2}(2i_l + f_l)
\]

\[
\geq \frac{s_l}{4\alpha \log n}
\]

\[
\geq \frac{s_{l+1} - 2\alpha k \log n}{4\alpha \log n}
\]

\[
= k.
\]

So it is sufficient to construct \( s_1, \ldots, s_l \) that satisfy (1). We define these stages inductively. For \( c = 0 \) we have \( s_0 = 0, i_0 = 0 \) and \( f_0 = 1 \), so (1) holds. Suppose we have determined some \( s_c \). If \( s_c > 4\alpha k \log n \), set \( l = c \) and we are done. Otherwise, we proceed as follows.

Let \( F \) be the set of frontier nodes at the beginning of stage \( s_c \), and let \( g \) be such that \( 2^{g-1} \leq |F| < 2^g \). For each \( j = 1, \ldots, g \), let \( Y_j \) be the set of nodes that received the message in stages \( s_c, s_c + 1, \ldots, s_c + a_j - 1 \) (but were inactive when stage \( s_c \) started).

We have two sub-cases:

\underline{Case 1}: There is \( j \) for which \( |Y_j| \geq 2^j \). In this case, take \( s_{c+1} = s_c + a_j \). At least \( |Y_j| \) new nodes received the message, so

\[
2i_{c+1} + f_{c+1} \geq 2i_c + f_c + |Y_j|
\]

\[
\geq \frac{s_c}{2\alpha \log n} + 2^j
\]

\[
\geq \frac{s_{c+1}}{2\alpha \log n}.
\]

\underline{Case 2}: For each \( j \) we have \( |Y_j| \leq 2^j \). We show that in this case all nodes in \( F \) will become inner after \( a_g \) stages.

Fix any node \( v \) that is inactive when stage \( s_c \) starts, and whose set \( X \) of in-neighbors in \( F \) is not empty. Pick \( j \) such that \( 2^{j-1} \leq |X| < 2^j \). Since \( |Y_j| \leq 2^j \), by property (\#), family \( S_j \) contains a set \( S_{j,i} \) that hits \( X \) and avoids \( Y_j \). This \( S_{j,i} \) will occur in one of the stages \( s_c, s_c + 1, \ldots, s_c + a_j - 1 \).

All in-neighbors of \( v \) are either in \( X \) or are inactive when stage \( s_c \) starts. When we use \( S_{j,i} \) for transmission then:

(i) exactly one in-neighbor of \( v \) in \( X \) will transmit because \( |S_{j,i} \cap X| = 1 \),

(ii) the nodes from \( Y_j \) will not interfere because \( S_{j,i} \cap Y_j = \emptyset \),
(iii) the nodes that were inactive at the beginning of stage $s_c$ and are not in $Y_j$ remain inactive when $S_{j,i}$ is issued, so they will not transmit.

Therefore $v$ will receive the message when $S_{j,i}$ is issued (unless it has already received it earlier). Since $v$ was an arbitrary inactive out-neighbor of $F$, we conclude that all nodes in $F$ will become inner after $a_g$ stages.

Take $s_{c+1} = s_c + a_g$. In this case $i_{c+1} \geq i_c + f_c$ and $f_c \geq 2^{a_g-1}$, so

$$2i_{c+1} + f_{c+1} \geq 2i_c + f_c + f_c$$

$$\geq \frac{s_c}{2\alpha \log n} + 2^{a_g-1}$$

$$\geq \frac{s_{c+1}}{2\alpha \log n}.$$  

We thus proved that (1) holds for $s_{c+1}$. Further, since $|F| \leq k$, we have $a_g = \alpha 2^{a_g} \log n \leq 2\alpha k \log n$, and thus in both sub-cases we have $s_{c+1} - s_c \leq 2\alpha k \log n$. The proof of the lemma is now complete.

4 Distributed Coupon Collection

In each phase of our algorithm we will attempt to distribute each message $m_v$ to some number of nodes, by performing a sequence of limited broadcasts. We need to achieve two goals. To obtain a good running time, the number of limited broadcasts must be small. Further, each message $m_v$ should participate in at least one limited broadcast, that is, $m_v$ must be in at least one $M_u$, for some $u$ that initiates a limited broadcast.

To choose the nodes $v$ for which we initiate a limited broadcast, we use randomization. The principle behind the random process we use is similar to that in the coupon collector’s problem. There are two differences though. First, each coupon may have several copies. Second, since we do not have enough time to coordinate the choices, we cannot guarantee that exactly one node will initiate broadcasting.

We think of $V$ as a set of $n$ bins and $M$ as a set of $n$ coupons. Each coupon has at least $k$ copies, each copy belonging to a different bin. $M_v$ is the set of coupons in bin $v$. Consider the following process: At each step, we open bins at random, by choosing each bin, independently, with probability $1/n$. If exactly one bin, say $v$, is opened, all coupons from $M_v$ are collected. If no bin is opened, or if two or more bins are opened, a failure occurs and no coupons are collected. How many steps do we need so that with high probability (a copy of) each coupon is collected?

The distributed coupon collection can be written in pseudocode as follows.
Procedure DistCouponColl($s$).

repeat $s$ times
  for each bin $v$ do
    with probability $1/n$ do open $v$
    else close $v$
  if exactly one bin $v$ is opened then
    collect all coupons from $M_v$

Lemma 2. Assume we have $n$ bins and $n$ coupons, and that each coupon has at least $k$ copies, each copy belonging to a different bin. Let $\delta$ be a given constant, $0 < \delta < 1$, and $s = (4n/k) \ln(n/\delta)$. Then, after performing DistCouponColl($s$), with probability at least $1 - \delta$, all coupons will be collected.

Proof. The lemma is trivially true for $n = 1$, so we can assume that $n \geq 2$. Let $\mathcal{X}_{m,j}$ be the event that coupon $m$ is collected at a given step $j$. Then $\Pr[\mathcal{X}_{m,j}]$ is the probability that one bin containing $m$ is opened and all other bins are closed.

For all $m$ and $j$ we have

$$\Pr[\mathcal{X}_{m,j}] \geq \frac{k}{n} \left(1 - \frac{1}{n}\right)^{n-1}$$

$$\geq \frac{k}{n} \left(1 - \frac{1}{n}\right)^n$$

$$\geq \frac{k}{4n},$$

where the last inequality follows from the fact that the sequence $(1 - 1/n)^n$ is monotonically increasing. The probability that some $m$ is not collected in $s$ steps is

$$\Pr \left[ \bigvee_{m \in M} \bigwedge_{j \leq s} \neg \mathcal{X}_{m,j} \right] \leq n \left(1 - \frac{k}{4n}\right)^s$$

$$\leq ne^{-sk/4n}$$

$$\leq \delta,$$

by the definition of $s$.

5 The Gossiping Algorithm

We now present our Monte Carlo algorithm for gossiping. Each node $v$ performs its version of the algorithm.
Algorithm RandGossip$_v(\epsilon)$.

\[ \delta \leftarrow \epsilon / \log n \]

for \( i = 0, 1, \ldots, \log n - 1 \) do

Phase$_i$:

\[ s_i \leftarrow (4n/2^i) \ln(n/\delta) \]

repeat \( s_i \) times

for each node \( v \) do (in parallel)

with probability \( 1/n \) do \( \text{active}_v \leftarrow \text{true} \)

else \( \text{active}_v \leftarrow \text{false} \)

\( \text{LtdBroadcast}_v(2^{i+1}) \)

Theorem 1. Let \( \epsilon, 0 < \epsilon < 1 \), be a given constant. With probability at least \( 1 - \epsilon \), Algorithm RandGossip$(\epsilon)$ completes gossiping in time \( O(n \log^3 n \log(n/\epsilon)) \).

Proof. In phase \( i \), the call to \( \text{LtdBroadcast}_v(2^{i+1}) \) costs \( O(2^i \log^2 n) \), so phase \( i \) costs \( O(s_i \cdot 2^i \log^2 n) = O(n \log^2 n \log(n/\delta)) \). Since we have \( \log n \) phases and \( \delta = \epsilon / \log n \), this implies the bound on the running time of the algorithm.

Initially, when phase 0 starts, each \( m_u \) is in one set \( M_v \), namely in \( M_u \). The algorithm attempts to maintain the invariant that after phase \( i \) each \( m_u \) is in at least \( 2^{i+1} \) sets \( M_v \). If this invariant is preserved at each phase, the gossiping will complete successfully, since after phase \( \log n - 1 \) each \( m_u \) will be in \( n \) sets \( M_v \). Thus it is sufficient to prove that the probability that the invariant fails in some phase is at most \( \epsilon \).

The process of distributing messages in a given phase \( i \) is equivalent to the distributed coupon collection problem described in the previous section, where we view each node as a bin, and active nodes correspond to open bins. Thus, by Lemma 2 the probability that the invariant fails in this phase, assuming that it has not failed in any previous phase, is at most \( \delta \). So, overall, the probability of failure in some phase is at most \( \log n \cdot \delta = \epsilon \).

To obtain a Las Vegas algorithm, run RandGossip$(\epsilon)$ with \( \epsilon = 1/n \). After the algorithm halts, each node that has not received \( n \) different initial messages, announces that a failure occurred. Since all these nodes send out the same message, this can be achieved with just one broadcast. If \( B \) is the running time of the broadcasting algorithm, after \( B \) steps each node knows whether the gossiping was successful or not. If the gossiping failed, we run a naive, deterministic RoundRobin algorithm that consists of \( n \) rounds, with each node transmitting once in each round. This will achieve gossiping in time \( O(n^2) \). Overall, the expected running time will be \( O((1 - 1/n)n \log^4 n + (1/n)n^2) = O(n \log^4 n) \). Concluding, we get the following theorem.

Theorem 2. There exists a randomized Las Vegas algorithm for gossiping with running time \( O(n \log^4 n) \).
6 Final Comments

Several open problems remain. The only known lower bounds for gossiping are those for broadcasting \[157\], a seemingly easier problem. The gap between lower and upper bounds is particularly wide in the deterministic case: between \(\Omega(n \log n)\) and \(O(n^{3/2} \log^2 n)\). Closing or at least reducing this gap is an interesting open problem.

Our algorithm is probably not optimal. One possible research direction is to investigate whether one can improve the running time of our algorithm by using the randomized broadcasting algorithm from \[3\] to perform limited broadcast. It is not quite clear whether the algorithm from \[3\] can be modified to satisfy Lemma 1. Further, the analysis of this modified algorithm will probably be much more complicated.

Improving the bounds on gossiping or broadcasting may resolve the question whether gossiping is harder than broadcasting in the radio network model (both in the randomized and the deterministic case). Our result implies that, at least in the randomized case, the difference is at most a poly-logarithmic factor.

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References

Deterministic Application of Grover’s Quantum Search Algorithm

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Abstract. Grover’s quantum search algorithm finds one of \( t \) solutions in \( N \) candidates by using \( (\pi/4)\sqrt{N/t} \) basic steps. It is, however, necessary to know the number \( t \) of solutions in advance for using the Grover’s algorithm directly. On the other hand, Boyer et al. proposed a randomized application of Grover’s algorithm, which runs, on average, in \( O(\sqrt{N/t}) \) basic steps (more precisely, \( (9/4)\sqrt{N/t} \) steps) without knowing \( t \) in advance. Here we show a simple (almost trivial) deterministic application of Grover’s algorithm also works and finds a solution in \( O(\sqrt{N/t}) \) basic steps (more precisely, \( (8\pi/3)\sqrt{N/t} \) steps) on average.

1 Introduction

Grover \[\text{Gro96, Gro97}\] proposed a quantum algorithm — Grover’s search algorithm — that solves the following general search problem much faster than any randomized/deterministic algorithm designed on classical computers.

Search Problem

**Given:** For any \( n > 0 \), an oracle Boolean function \( f \) on the set \( \{0, 1\}^n \) of binary sequences of length \( n \).

**Question:** Find some sequence \( x \in \{0, 1\}^n \) such that \( f(x) = 1 \).

Remark. In general, a solution, i.e., a sequence \( x \in \{0, 1\}^n \) satisfying \( f(x) = 1 \), is not unique, in which case it is sufficient to output *any one* of such solutions.

For any given \( n \), since there are \( N = 2^n \) binary strings in \( \{0, 1\}^n \), it is (almost obvious) that any algorithm solving the above problem needs \( N \) steps to find the desired sequence. Surprisingly, though, Grover’s search algorithm finds the desired sequence in \( O(\sqrt{N}) \) quantum computation steps, where each quantum step (which we refer to as G-steps) can be implemented by some \( \text{poly}(n) \) number of basic quantum gates. More precisely, when there are \( t \) solutions, i.e., \( t \) binary sequences satisfying \( f \), Grover’s algorithm finds some of them in \( O(\sqrt{N/t}) \) G-steps. Note, however, that one needs to know the number \( t \) in advance in order to achieve this better bound; but \( t \), the number of solutions, is usually unknown in advance. For solving this problem, Boyer et al. \[\text{BBHT96}\] proposed an algorithm that applies Grover’s algorithm with randomly chosen parameters, that...
is, a randomized application of Grover’s algorithm, which runs, on average, in $O(\sqrt{N/t})$ G-steps without knowing $t$ in advance. Here we show that a simple (in fact, almost trivial) deterministic application of Grover’s algorithm also works and finds a solution in $O(\sqrt{N/t})$ G-steps on average without knowing $t$ in advance. (Notice also here that any classical computation including randomized one can be simulated by some quantum computation. In fact, for any classical algorithm using any quantum algorithm(s), one can design a single quantum Turing machine that executes the whole algorithm. In this sense, there is no merit of removing randomness from the application algorithm. Yet, the authors think that the simpler algorithm is the better.)

**Grover’s Algorithm and Its Randomized Application**

Let us discuss our problem specifically. We start with recalling Grover’s algorithm and some basic facts about the algorithm. Notions and notations for quantum computation we use here are standard; see, e.g., [BV97, Gru99, Hos99, For00].

Consider any $n > 0$ and any oracle function $f$ on $\{0, 1\}^n$ for specifying the above general search problem, and let us fix them in the following discussion. That is, our task is to find a sequence $x \in \{0, 1\}^n$ that satisfies $f(x) = 1$. A sequence $x \in \{0, 1\}^n$ satisfying $f(x) = 1$ is simply called a solution. Let $N = 2^n$, i.e., the total number of sequences in $\{0, 1\}^n$, and let $t$ denote the number of all solutions among $N$ candidates.

In Grover’s algorithm, each sequence $x \in \{0, 1\}^n$ corresponds to a quantum base state $|x\rangle$ consisting of $n$ qubits. (In the following, we identify an $n$ qubit base state with the corresponding $n$ bit binary sequence.) The main ingredients of Grover’s algorithm are the following three unitary transformations on $n$ qubit states.

**Walsh-Hadamard transformation:**

$$W : |x\rangle \mapsto \frac{1}{\sqrt{N}} \sum_{y \in \{0, 1\}^n} (-1)^{x \cdot y} |y\rangle$$

(Here $x \cdot y$ denotes the bit-wise inner product.)

**Sign flipping on $f$:**

$$S_f : |x\rangle \mapsto (-1)^{f(x)} |x\rangle$$

**Sign flipping on 0:**

$$S_0 : |0\rangle \mapsto - |0\rangle \text{ and } |x\rangle \mapsto |x\rangle \text{ (if } x \neq 0^n \text{)}$$

By using these transformations, one G-step is defined as the following unitary transformation $U$.

$$U = -WS_0WS_f$$

That is, *one G-step* is to apply this $U$ to a current state. Grover’s algorithm is to apply $U$ for some appropriate number of times to the following initial state $\phi_0$. 
\[ |\psi_0\rangle = \sum_{a \in \{0,1\}^n} \frac{1}{\sqrt{N}} |a\rangle. \]

Formally, by the \textit{j G-step execution of Grover’s algorithm} (or more simply \(G(j)\)) we mean to apply \(UUU \cdots U\) to \(\phi_0\) and then observe the obtained state. (We assume that the observation is made so that some \(n\) qubit base state (i.e., \(n\) bit sequence) is observed with the probability that is the square of its amplitude.)

For justifying this procedure, the following property of \(U\) plays a key role [Gro96, Gro97].

\textbf{Lemma 1.} Consider the quantum state obtained by applying \(U\) to \(\phi_0\) for \(j\) times. Then each solution state, i.e., a base state corresponding to a solution, has the same amplitude (which we denote \(a_j\)) while the other base state also has the same amplitude (which we denote \(b_j\)). Furthermore, by using \(\theta\) that satisfies \(\sin \theta = \sqrt{t/N}\), these amplitudes are stated as follows.

\[ a_j = \frac{1}{\sqrt{t}} \sin((2j + 1)\theta), \]
\[ b_j = \frac{1}{\sqrt{N - t}} \cos((2j + 1)\theta). \]

The angle \((2j + 1)\theta\) determines the amplitude \(a_j\). In the following we call this angle \textit{the angle (of a solution) after applying \(U\) for \(j\) times (to the initial state)} (or more simply, \textit{the angle after executing \(G(j)\)}). Note that after applying \(U\) for \(\lceil(\pi/4)\theta\rceil\) times, the angle gets close to \(\pi/2\); hence, the amplitude of each solution is close to \(1/\sqrt{t}\), which means that the total probability that solution states are observed is close to 1. Note also that \((\pi/4)\theta\) is approximately \((\pi/4)\sqrt{N/t}\) by using the approximation \(\theta \approx \sin \theta = \sqrt{t/N}\). This argument leads us to the following theorem of Grover [Gro96, Gro97].

\textbf{Theorem 1.} Define \(m_0\) by

\[ m_0 = \left\lfloor \frac{\pi}{4\theta} \right\rfloor \left( \approx \frac{\pi}{4} \sqrt{\frac{N}{t}} \right). \]

Then a state observed by the execution of \(G(m_0)\) is one of the solutions with probability approximately \(1 - 1/N (\approx 1)\).

Therefore, if we can compute \(m_0\), we would simply execute \(G(m_0)\) to get some solution. It is, however, not so easy to compute it because \(t\) is usually unknown in advance. Note that we cannot simply execute \(G(m)\) with some \(m > m_0\); in this case, the probability that some solution is observed could become much smaller. In order to solve this problem, Boyer et al. [BBHT96] proposed the following randomized algorithm.

It is clear that this algorithm finds a solution. On the other hand, we can show the following time bound.
Algorithm Randomized-Grover
\[
\lambda \leftarrow \frac{6}{5}; \ i \leftarrow 0;
\]
for \( i \leftarrow 1 \) to \( \infty \) do
\[
m \leftarrow \lambda^i;
\]
select \( j \), \( 0 \leq j \leq m - 1 \), uniformly at random;
execute \( G(j) \) and let \( x \) be the observed state;
if \( f(x) = 1 \) then output \( x \) and halt;
end-for;

Fig. 1. A randomized application of Grover’s algorithm

Theorem 2. The average number of \( G \)-steps executed in the above algorithm is at most \((9/4) \sqrt{N/t}\).

Remark. In the proof of this theorem and in the following discussion in general, we often assume that \( t/N \) is small enough, say, \( t/N < 1/100 \) and use some approximations like \( \sqrt{t/N} \approx \sin \sqrt{t/N} \) under this assumption. Thus, our bounds may not be precise; but even so, it is easy to see that the bounds shown here certainly hold by changing constant factors. We leave the tedious precise calculation to the interest reader. Also though our arguments need an assumption such as \( t < N/100 \), this assumption is not essential for designing a fast search algorithm. In fact, if \( t \geq N/100 \), then we can search for a solution by simply picking any \( n \) bit sequence randomly, say, 1,000 times.

2 Deterministic Application of Grover’s Algorithm

Here we show that a simple and deterministic execution of Grover’s algorithm still yields a similar result. More specifically, we consider the algorithm given in Figure 2

Algorithm Deterministic-Grover
\[
\%
\text{The parameter } k \text{ is fixed to 2.}
\]
for \( i \leftarrow 1 \) to \( \infty \) do
\[
m \leftarrow 2^i;
\]
execute \( G(m) \) for \( k \) times and let \( x_1, \ldots, x_k \) be the observed states;
if \( f(x_u) = 1 \) for some \( u \), \( 1 \leq u \leq k \) then output \( x_u \) and halt;
end-for;

Fig. 2. A deterministic application of Grover’s algorithm

Again for this algorithm, we can show that its average running time is \( O(\sqrt{N/t}) \). In the following we will prove this fact after preparing some tools.

First let us see that at some point of the for-iteration of the algorithm, the angle after executing \( G(m) \) in this iteration gets reasonably close to \( \pi/2 \).
Claim. There exists some integer \( i_0 > 0 \) such that
\[
\frac{\pi}{3} \leq (2 \cdot 2^{i_0} + 1)\theta \leq \frac{2\pi}{3}.
\]
Furthermore, since \( \theta \approx \sqrt{t/N} \), we have \( 2^{i_0} < (\pi/3)\sqrt{N/t} \).

Proof. Clearly, \((2 \cdot 2^{i} + 1)\theta \approx (2 \cdot 2^{i} + 1)\sqrt{t/N} > \pi/3 \) for some \( i > 0 \). Let \( i \) be the largest integer such that \((2 \cdot 2^{i} + 1)\theta < \pi/3 \) holds. Then we have
\[
(2 \cdot 2^{i+1} + 1)\theta < 2(2 \cdot 2^{i} + 1)\theta < \frac{2\pi}{3}.
\]
Therefore, the claim holds with \( i_0 = i + 1 \).

Below we will keep using \( i_0 \) to denote the one satisfying this claim. Also let \( W \) denote the above range of angles; that is, \( W = [\pi/3, 2\pi/3] \). As stated in the above claim, the angle at the \( i_0 \)th for-iteration gets into \( W \); then the probability that a solution is found at this for-iteration is, as we will see below, more than \( 1 - (1/4)^k \). In general, for any \( i \), let us simply call the angle \((2 \cdot 2^{i} + 1)\theta \) the angle at the \( i \)th for-iteration. Recall that the success probability is defined by \( \sin^2 \) of angles; hence, we had better discuss angles with their residues under modulo \( \pi \). Thus, in the following, by “an angle” we mean the residue of a real angle under modulo \( \pi \). For any for-iteration, we have the following relation between the angle and the failure probability at this for-iteration.

Claim. For any \( i > 0 \), let \( \omega \) be the angle at the \( i \)th for-iteration. Let \( \alpha = |\pi/2 - \omega| \) and \( \delta = \sin^2 \alpha \). Then the probability that no solution is found at this for-iteration is \( \delta^k \). (Note that if \( \omega \in W \), then we have \( \delta \leq 1/4 \).)

Proof. Since the angle at the \( i \)th for-iteration is \( \omega \), the probability that some solution is found by one execution of \( G(m) \) (where \( m = 2^i \)) is
\[
t \cdot \left( \frac{1}{\sqrt{t}} \sin \omega \right)^2 = \sin^2 \omega.
\]
Hence, the probability that no solution is found at this iteration is
\[
(1 - \sin^2 \omega)^k = (\cos^2 \omega)^k = (\sin^2 |\pi/2 - \omega|)^k = (\sin^2 \alpha)^k = \delta^k.
\]
It would be nice if we could argue that the angle belongs to \( W \) reasonably often. Unfortunately, however, if the angle (at some for-iteration) gets very close to \( \pi/2 \), then it takes rather long time to have an angle in \( W \) again. Note, on the other hand, that if the angle is close to \( \pi/2 \), then the probability of finding a solution at this iteration is close to 1. The crucial point of our analysis is to estimate this trade-off.

For simplifying our discussion, instead of an angle \( \omega \), we will argue by using \( \alpha = |\pi/2 - \omega| \), which we call a \textit{co-angle}. For example, the above claim shows that if the co-angle (at some for-iteration) gets less than \( \pi/6 \), then the error
probability at this iteration becomes less than \((1/4)^k\). Let \(A\) denote this good co-angle range; that is, \(A = \{\alpha : 0 \leq \alpha \leq \pi/6\}\).

For our analysis, we estimate the number of for-iterations until having a co-angle in \(A\) again.

**Claim.** Consider any \(i\)th for-iteration with the co-angle \(\alpha \in A\). (We assume that \(\alpha > 0\).) Define \(h = \log(\pi/(3\alpha))\). Then there exists some \(\overline{h}, 1 \leq \overline{h} \leq h\), such that the co-angle at the \((i + \overline{h})\)th for-iteration gets into \(A\) again.

**Remark.** For simplifying our discussion, the argument here is a bit unprecise. More specifically, it may be the case that the co-angle at the \((i+\overline{h})\)th for-iteration is slightly larger than \(\pi/3\); but the difference is at most \(+3\theta\), which should be small enough under our assumption on \(t/N\).

**Proof.** Here we show that the angle at the \((i + \overline{h})\)th for-iteration belongs to the range \(W = [\pi/3, 2\pi/3]\).

Since the claim clearly holds if \(\alpha = \pi/6\), we may assume that \(\alpha < \pi/6\), which implies that \(h \geq 2\). Recall that the definitions of angle and co-angle; that is, we have \((\omega =) \pi/2 - \alpha = (2 \cdot 2^i + 1)\theta\), where \(\omega\) is the angle at the \(i\)th for-iteration. (Here we consider only the case that \(\omega < \pi/2\). The case that \(\omega > \pi/2\) can be treated in a similar way.) Then for any \(h', 2 \leq h' \leq h\), we can estimate the angle \(\omega'\) at the \((i + h')\)th for-iteration as follows.

\[
\omega' = (2 \cdot 2^{i+h'} + 1)\theta \mod \pi
\]
\[
= 2^{h'} \cdot ((2 \cdot 2^i + 1)\theta) - (2^{h'} - 1)\theta \mod \pi
\]
\[
= 2^{h'} \cdot \left(\frac{\pi}{2} - \alpha\right) - (2^{h'} - 1)\theta \mod \pi
\]
\[
= -(2^{h'} \alpha + (2^{h'} - 1)\theta - m\pi) \mod \pi.
\]

Here \(m\) is some nonnegative integer such that \(0 \leq 2^{h'} \alpha + (2^{h'} - 1)\theta - m\pi < \pi\). Since \(\sin \omega' = \sin(\pi - \omega')\), we can discuss by using \(\varpi' = 2^{h'} \alpha + (2^{h'} - 1)\theta - m\pi\) instead of the above \(\omega'\).

Now we consider the first \(h' \leq h\) such that \(2^{h'} \alpha + (2^{h'} - 1)\theta \geq \pi/3\). Such \(h'\) exists because we have \(2^h \alpha \geq \pi/3\) from our choice of \(h\). Furthermore, we may assume that \(h' \geq 2\), because in the special case that \(h' = 1\), we can easily show that \(\overline{h} = 2\) satisfies the claim. Also notice that \(\varpi' = 2^{h'} \alpha + (2^{h'} - 1)\theta\); that is, \(m = 0\) for this \(h'\).

Clearly, we are done if \(\varpi' \leq 2\pi/3\). But if otherwise, i.e., if \(2^{h'} \alpha + (2^{h'} - 1)\theta > 2\pi/3\), then the angle \(\varpi'' = 2^{h'-1} \alpha + (2^{h'-1} - 1)\theta \approx (2^{h'} \alpha + (2^{h'} - 1)\theta)/2\) should be \(\geq \pi/3\). On the other hand, since \(2^{h'} \alpha + (2^{h'} - 1)\theta < \pi\), we have \(\varpi' = 2^{h'-1} \alpha + (2^{h'-1} - 1)\theta < \pi/2\). Therefore, the claim is satisfied with either \(\overline{h} = h'\) or \(\overline{h} = h' - 1\).

Next we divide the co-angle range \(A\). Below let \(j\) be any nonnegative integer. Let \(\delta_j = 2^{-(j+2)}\), and let \(\tilde{\alpha}_j\) be the co-angle \(< \pi/2\) such that \(\delta_j = \sin^2 \tilde{\alpha}_j\). (For example, we have \(\delta_0 = 1/4\) and \(\tilde{\alpha}_0 = \pi/6\).) Then define \(A_j = \{\alpha : \tilde{\alpha}_{j+1} < \alpha \leq \tilde{\alpha}_j\}\). Also we define \(h_j\) as follows.
Since $\hat{\alpha}_j$ gets very small, we may assume that $\sin \hat{\alpha}_j \approx \hat{\alpha}_j$. In particular, we may assume that $\hat{\alpha}_{j+1} \geq \hat{\alpha}_j/2$. (In fact, we would have $\hat{\alpha}_{j+1} \geq \hat{\alpha}_j/\sqrt{2}$ for large $j$.) Hence, we have $h_{j+1} \leq h_j + 1$.

Here let us summarize our discussion. Consider any $i$th for-iteration, and suppose that the co-angle at this iteration belongs to $A_j$. Then it follows from the above claims that (i) the probability that the algorithm does not halt at this for-iteration is $\leq \delta_j^k (= 2^{-(i+2)k})$, and (ii) the co-angle gets into $A$ again before the $(i + h_j)$th for-iteration.

Now we are ready to prove our theorem.

**Theorem 3.** The average number of G-steps executed in Deterministic Grover (the algorithm of Figure 2) is $(8\pi/3)\sqrt{N/\ell}$.

**Proof.** Recall that the co-angle belongs to $A$ at the $i_0$th for-iteration. We may assume that it is indeed the first for-iteration having a co-angle in $A$. In general, we use $i_1, i_2, ..., \text{ to denote the first, the second, } \text{ ... for-iteration (after the } i_0 \text{th for-iteration) where the co-angle belongs to } A. \text{ Also let } \alpha_0, \alpha_1, \text{ ... denote the co-angles at the } i_0 \text{th, } i_1 \text{th, } \text{ ... iterations.}

To bound the expected number of executed G-steps, we may assume that the execution of the algorithm reaches to the $i_0$th for-iteration. Thus, we first estimate the number of G-steps executed by the end of the $i_0$th for-iteration, which is bounded as follows.

$$k \cdot (1 + 2 + 2^2 + \cdots + 2^{i_0}) = k \cdot (2^{i_0+1} - 1) < \frac{2k\pi}{3} \sqrt{\frac{N}{\ell}}. \quad (1)$$

Next we analyze the expected number $E_0$ of G-steps executed after the $i_0$th for-iteration. (Precisely speaking, $E_0$ that we will analyze below is the expected number $E_0$ of G-steps executed after the $i_0$th for-iteration under the condition that the $i_0$th for-iteration is executed.) Although we know that the co-angle $\alpha_0$ at the $i_0$th for-iteration is in $A$, we do not know which $A_j$ it belongs to. But for each $j \geq 0$, by assuming that $\alpha \in A_j$, we can estimate the failure probability and the length to the $i_1$th for-iteration. Thus, we can bound $E_0$ in the following way. (Here we use $E_1$ to denote the amount corresponding to $E_0$; that is, $E_1$ is the expected number of G-steps executed after the $i_1$th for-iteration under the condition that the $i_1$th for-iteration is executed.)

$$E_0 \leq \sum_{j \geq 0} \Pr\{ \text{ no solution is found at the } i_0 \text{th for-iteration } | \alpha_0 \in A_j \} \times (\# \text{ of G-steps executed during } ((i_0 + 1) \sim i_1) \text{th for-iteration} + E_1) \leq \sum_{j \geq 0} \delta_j^k \cdot (k \cdot (2^{i_0+1} + 2^{i_0+2} + \cdots + 2^{i_0+h_j}) + E_1)$$
Here by using the fact that \( h_j \leq h_0 + j \) and \( h_0 \leq 2 \), and our choice of \( \delta_j \) and \( k \) (i.e., \( \delta_j = 2^{-j+2} \) and \( k = 2 \)), we can bound the last expression as follows.

\[
E_0 \leq k2^{i_0} \cdot \left( \sum_{j \geq 0} 2^{-j} \cdot 2^{h_0 + j} \right) + \sum_{j \geq 0} \delta_j \cdot E_1 = k2^{i_0} + \sum_{j \geq 0} \delta_j \cdot E_1. \tag{2}
\]

At this point, let us see how \( E_1 \) is estimated. Notice that \( E_1 \) depends on the value of \( i_1 \). Otherwise, the calculation is quite similar to the one for \( E_0 \). (Here we use \( E_2 \) denoting the amount corresponding to \( E_1 \) for \( i_2 \).)

\[
E_1 \leq \sum_{j \geq 0} \Pr\{ \text{no solution is found at the } i_1 \text{th for-iteration } | \alpha_1 \in A_j \} \times (\# \text{ of G-steps executed during } ((i_1 + 1) \sim i_2)\text{th for-iteration} + E_2)
\]

\[
\leq \sum_{j \geq 0} \delta_j \cdot \left( k \cdot (2^{i_1 + 1} + 2^{i_1+2} + \cdots + 2^{i_1+h_j}) + E_2 \right)
\]

\[
\leq k2^{i_1} \cdot \left( \sum_{j \geq 0} \delta_j \cdot 2^{h_j+1} \right) + \sum_{j \geq 0} \delta_j \cdot E_2 \leq k2^{i_1} + \sum_{j \geq 0} \delta_j \cdot E_2. \tag{3}
\]

We substitute \( E_1 \) of (2) by (3). Here notice that \( i_1 \) depends on the choice of \( \alpha_0 \); in fact, if \( \alpha_0 \in A_j \), then it is bounded by \( i_0 + h_j \) (\( \leq i_0 + h_0 + j \leq i_0 + j + 2 \)). Thus, we have the following bound.

\[
E_0 \leq k2^{i_0} + \sum_{j_1 \geq 0} \delta_{j_1} \cdot \left( k2^{i_1} + \sum_{j_2 \geq 0} \delta_{j_2} \cdot E_2 \right)
\]

\[
\leq k2^{i_0} + \sum_{j_1 \geq 0} 2^{-2(j_1+2)} \cdot k2^{i_0+j_1+2} + \sum_{j_1 \geq 0} \delta_{j_1} \cdot \left( \sum_{j_2 \geq 0} \delta_{j_2} \cdot E_2 \right)
\]

\[
= k2^{i_0} + k2^{i_0} \sum_{j_1 \geq 0} 2^{-(j_1+2)} + \sum_{j_1 \geq 0} \delta_{j_1} \cdot \left( \sum_{j_2 \geq 0} \delta_{j_2} \cdot E_2 \right)
\]

\[
= k2^{i_0} + \frac{1}{2} \cdot k2^{i_0} + \sum_{j_1 \geq 0} \delta_{j_1} \cdot \left( \sum_{j_2 \geq 0} \delta_{j_2} \cdot E_2 \right).
\]
We can similarly expand $E_2$, $E_3$, ..., thereby deriving the following bound.

$$E_0 \leq k2^{i_0} + 2^{-1} \cdot k2^{i_0} + 2^{-2} \cdot k2^{i_0} + \cdots = 2k2^{i_0} \leq \frac{2k\pi}{3} \sqrt{\frac{N}{t}}. \quad (4)$$

Now the bound of the theorem is immediate from (1) and (4), and our choice of $k$.

References


Random Instance Generation for MAX 3SAT
(Extended Abstract)

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1 Introduction

MAX SAT is one of famous combinatorial optimization problems stated as follows: given a multiset of clauses, find an assignment that maximizes the number of satisfied clauses (that is equivalent to finding an assignment that minimizes the number of unsatisfied clauses). MAX 3SAT is restricted version of MAX SAT, that is, its input is restricted to a multiset of 3-clauses, i.e., each clause contains exactly 3 literals whose underlying variables are distinct each other. Since these problems are not only NP-hard problem, but MAX SNP-complete problem, there is no polynomial time approximation algorithm whose approximation ratio is close to 1 unless P = NP. Furthermore, Håstad showed that it is NP-hard to approximate MAX 3SAT within $8/7 - \varepsilon$ for any $\varepsilon > 0$ [5]. In spite of these negative results, many polynomial time approximation algorithms with proven approximation ratio for MAX SAT have been proposed [3,4,7].

To evaluate these approximation and exact algorithms experimentally, test instances are necessary. While there are some benchmark problems for MAX SAT, we still do not have enough number of problems, and we would like to have sure way generating nontrivial test instance systematically. In particular, it would be nice if we have an efficient test instance generators for MAX SAT.

Test instance generator have been studied for NP problems (see, e.g. [1, 2, 6, 8]). On the other hand, for optimization problems like MAX SAT, no generator, as far as the author knows, has been proposed in the literature.

When one evaluates an approximation algorithm, test instances should have their optimal solution otherwise we can not determine how good approximated solution are obtained. In spite of necessity of optimal solution, it seems hard to check whether given a (candidate of) optimal solution is really optimal. In fact, in the case of MAX SAT, it is co-NP complete to check the optimality of a given solution.

Since the hardness of generating a test instance with its optimal solution seems unavoidable, we consider two alternative approaches. The first one is to estimate (analytically and experimentally) the expected number $E[s(F)]$ of optimally satisfied clauses of $F$, when MAX SAT instance $F$ is generated under a certain distribution. Suppose that we can estimate $E[s(F)]$ with reasonable precision. Then even if we can not supply optimal solutions, we could still test
the average performance of a given MAX SAT algorithm. That is to say, we just run it on generated instances and compare the average number of clauses satisfied by obtained solution with \( E[s(F)] \). The second approach is to design a test instance generator that provides both an instance and its candidate of optimal solution that is really optimal with high probability.

In this paper, we obtain some basic results for these two approaches. For the first one, we analyze the minimum number \( u(F) \) of unsatisfiable clauses of given formula \( F \). We obtain two lower bounds on its expectation \( E[u(F)] \), \( \Omega(m/\ln m) \) and \( m/8 - O(\sqrt{mn}) \), for random 3CNF formula with \( m \) clauses over \( n \) variables. We also analyze the expectation experimentally. For the second approach, on the other hand, we propose a simple random instance generator that, for given \( u \), generates an instance and an assignment (i.e., a candidate for the optimal solution) that leaves just \( u \) clauses unsatisfied. Furthermore, we prove that the generated solution is indeed optimal with high probability if the number of clauses is \( \Omega(n \ln n) + \Omega(n \sqrt{u \ln n}) \).

Hereafter, we describe basic notations. Let \( V = \{v_1, \ldots, v_n\} \) be propositional variables. We represent true and false by 1 and 0 respectively. An assignment \( t \) is a mapping from variables to Boolean values, \( v_1, \ldots, v_n \rightarrow \{0, 1\}^n \). We denote \( T_d \) a set of assignments whose Hamming distance from \( 1^n \) is \( d \). A literal is a propositional variable \( v \) or its negation \( \overline{v} \). We use a complement-free set of exactly 3 literals to denote a 3-clause, a disjunction of exactly 3 literals. Let 3CNF formula \( F \) be a multiset of 3-clauses, and we denote \( |F| \) by \( m \). We say \( F \) is a random 3CNF formula with \( m \) clauses over \( n \) variables if each clause in \( F \) is chosen with replacement from possible \( 8 \binom{n}{3} \) clauses over \( n \) variables uniformly at random. For any \( F \) and \( t \), we denote the number of unsatisfied clauses in \( F \) at an assignment \( t \) by \( u(F, t) \). Furthermore, \( u(F) \) is the minimum number of unsatisfied clauses in \( F \) for all assignments, that is, \( u(F) = \min_{t \in \{0, 1\}^n} u(F, t) \). We say \( u(F) \) is the optimal value of \( F \).

In this abstract, we omitted proofs and some explanations. See research report in http://www.is.titech.ac.jp/.

2 The Average Optimal Value for Random 3CNF Formula

We consider the expected number \( E[u(F)] \) of unsatisfiable clauses on the optimal solution. There exists an obvious upper bound \( m/8 \) since it is the upper bound of \( u(F) \). Here we prove two lower bounds of \( E[u(F)] \).

**Theorem 1.** For any \( m \geq n \ln 2 / \ln(8/7) \approx 5.19n \), the expected number of unsatisfiable clauses on optimal assignment satisfies the following bound,

\[
E[u(F)] \geq \frac{m \ln 8/7 - n \ln 2 + \ln(m/7 - 1)}{\ln m / 7} (1 - 7/m).
\]

The sketch of the proof is as follows. For any assignment \( t \), the probability \( \Pr[u(F, t) = v] \) that the number of unsatisfiable clauses in \( F \) on \( t \) is exactly \( v \) is estimated as \( \binom{m}{v} \left( \frac{1}{8} \right)^v \left( \frac{7}{8} \right)^{m-v} \). By using this estimation, we can bound
the probability that the number of unsatisfiable clauses in $F$ on its optimal assignment is larger than $v$. Finally, we use Markov’s inequality.

In the Theorem 1, rounding $\binom{m}{v}$ by $m^v$ causes the denominator of upper bound, $\ln m/7$. Thus we need to estimate the probability that $u(F)$ is larger than $v$ more tightly. The probability $\Pr[u(F) \geq v]$ includes partial sum of binomial distribution. For sufficiently large $m$, we can approximate binomial distribution to normal distribution by using central limit theorem. This observation leads us to the next result.

**Theorem 2.** For any $m \geq 7n \ln 2 + \sqrt{7n \ln 2} - 16\sqrt{7n \ln 2} - 8 \approx 9.7n - 8$, the expected number $E[u(F)]$ of optimally unsatisfiable clauses satisfies following bound,

$$E[u(F)] \geq \frac{1}{8} \left( m - \sqrt{14mn \ln 2} + 8 \right) \left( 1 - \frac{4}{\sqrt{7 \pi nm \ln 2}} \right).$$

We compare theoretical lower bounds with the average number of unsatisfiable clauses on optimal assignment that we obtained through some experiments. Now we describe the setup used in the experiments. We fixed the number of variables to 50, and we selected the numbers of clauses $m$ from 250 to 1000 at intervals of 50 (we skipped 900). For each $m$, we checked the average optimal value to 100 (50, for over 600 clauses) random 3CNF.

We show results in Fig. 1(a), along the horizontal axis we give the number of clauses $m$, and along the vertical axis we give the average number of unsatisfiable clauses on optimal. Big black dots are results of experiments, and dotted line and solid line are lower bounds of Theorem 1 and Theorem 2 resp. Furthermore, white circle are values obtained by calculating the partial sum of binomial distribution, $\arg \max_v \left( v \left( 1 - \sum_{j=0}^v \binom{m}{v} (1/8)^v (7/8)^{m-v} \right) \right)$. With this graph, it seems that behavior of experimental results is similar to that of lower bound of Theorem 2.

In Fig. 1(b), we show the ratio of difference between the experimental and the theoretical value to $m$. Black (white, resp.) dots are the ratio of the difference between the experimental value and the lower bound of Theorem 2 (the value that obtained by calculating the binomial distribution, resp.). We can see that the difference is as much as 0.02$m$. Thus we assume that the relation of $E[u(F)]$
with \( m \) forms \( (1/8-c_1)m-\sqrt{c_2m}+c_3 \), and get \((1/8-0.00620)m-\sqrt{1.13m}+10.9\) (dashed line in Fig. 1(a)) by the least squares.

3 Random Instance Generation for MAX 3SAT

Our final goal is to generate MAX SAT instance with its optimal solution at random. When evaluating MAX SAT solver experimentally, the optimal solution is necessary since it is hard to determine whether obtained solutions are correct. On the contrary, it is also hard to check a (candidate of) optimal solution generated with an instance. Hence, we relax the condition of instance generator such that the candidate of optimal solution is correct with high probability, that is, for given \( n \) and \( u \), our generator generates appropriate size MAX SAT instance over \( n \) variables and an assignment (for candidate of optimal solution) that unsatisfies exactly \( u \) clauses.

Our basic idea is as follows. Let \( F \) and \( \hat{t} \) be the output instance and assignment of our instance generator for given \( n \) and \( u \). We assume that \( 1^n \) is chosen as \( \hat{t} \) w.l.o.g. Of course, \( F \) shall have exactly \( u \) clauses that are not satisfied by \( \hat{t} \). These clauses shall be the form of \( \{ \overline{v_i}, v_j, \overline{v_k} \} \). There exist some assignments in \( T_1 \) that satisfy some of these \( u \) clauses while the number of unsatisfied clauses by \( 1^n \) is exactly \( u \). Therefore, for each of these \( u \) clauses, we add three more clauses forms of \( \{ v_i, \overline{v_i}, \overline{v_j} \}, \{ v_j, \overline{v_i}, \overline{v_j} \} \) and \( \{ v_k, \overline{v_k}, \overline{v_j} \} \), respectively. Now each assignment in \( T_1 \) has exactly \( u \) unsatisfied clauses. Then we add appropriate number of clauses that satisfied by \( \hat{t} \). Intuitively, \( \hat{t} \) becomes one of optimal solutions of \( F \) with high probability when enough large number of clauses are added.

Now we introduce our instance generation algorithm formally. First, we define some sets of clauses \( C_{9-}, C_{9+}, C_{5}^{i}, C_{3-}, \text{ and } C_{3+} \) as follows:

\[
C_{9-} = \{ \{ v_{i1}, \overline{v}_{i2}, \overline{v}_{i3} \}, \{ v_{i1}, \overline{v}_{i4}, v_{i5} \}, \{ v_{i2}, \overline{v}_{i6}, \overline{v}_{i7} \}, \{ v_{i1}, v_{i8}, \overline{v}_{i5} \} \},
\]

\[
C_{9+} = \{ \{ v_{i1}, v_{i2}, v_{i3} \}, \{ \overline{v}_{i1}, v_{i4}, \overline{v}_{i5} \}, \{ v_{i2}, v_{i6}, \overline{v}_{i7} \}, \{ v_{i1}, v_{i8}, v_{i9} \} \},
\]

\[
C_{5}^{i} = \{ \{ v_i, \overline{v}_{i1}, \overline{v}_{i2} \}, \{ \overline{v}_{i}, v_{i3}, v_{i4} \} \},
\]

\[
C_{3-} = \{ v_{i1}, \overline{v}_{i2}, \overline{v}_{i3} \}, \text{ and } C_{3+} = \{ \overline{v}_{i1}, v_{i2}, v_{i3} \}.
\]

Hereby we show a pseudo code of our algorithm in Fig. 2. In this algorithm, \( u \) many sets in the form of \( C_{9-} \) (\( C_{9+} \), resp.) guarantee that an assignment \( 1^n \) (\( 0^n \), resp.) does not satisfy \( u \) clauses exactly and each assignment in \( T_1 \) (\( T_{n-1} \), resp.) does not satisfy at least \( u \) clauses. On the other hand, there will exist some assignments that does not satisfy less than \( u \) clauses. In such case, \( \hat{t} \) is no longer the optimal solution of \( F \). To avoid these assignments, we add \( r \) clauses satisfied by both of \( 1^n \) and \( 0^n \) (\( C_{3-} \) and \( C_{3+} \)). We also add \( b \) many sets of \( \bigcup_{i=1}^{n} C_{5}^{i} \) to balance the number of occurrence of each literal. Intuitively, the larger \( r \) makes the higher probability that \( \hat{t} \) is the optimal solution of \( F \). Hence, we analyze the condition of \( r \) such that \( \hat{t} \) is the optimal solution of generated instance.

Let random variable \( X_t \) be the number of clause sets forms of \( C_{9-}, C_{9+}, C_{5}^{i}, C_{3-}, \text{ and } C_{3+} \) that contain at least one clause not satisfied \( t \). It is clear that
\textbf{input:} the number of variables $n$, the desirable optimal value $u$, the number of clause sets $\bigcup_{i=1}^{k}C_{5}^{i}$, $b$, the number of random clauses $r$.

\begin{align*}
\text{begin} \\
\text{let } F = \emptyset; \\
\text{repeat } u \text{ times do} \\
\hspace{1em} \text{choose 9 distinct variables } v_{r1}, \ldots, v_{r9} \text{ and add } C_{9-}(v_{r1}, \ldots, v_{r9}) \text{ to } F; \\
\hspace{1em} \text{choose 9 distinct variables } v_{r1}, \ldots, v_{r9} \text{ and add } C_{9+}(v_{r1}, \ldots, v_{r9}) \text{ to } F; \\
\hspace{1em} \text{repeat } b \text{ times do} \\
\hspace{2em} \text{for } i = 1 \text{ to } n \text{ do} \\
\hspace{3em} \text{choose 4 distinct variables } v_{r1}, \ldots, v_{r4} \in V \setminus \{v_{1}\} \text{ and add } C_{5}^{i}(v_{r1}, \ldots, v_{r4}) \text{ to } F; \\
\hspace{1em} \text{repeat } r \text{ times do} \\
\hspace{2em} \text{choose 3 distinct variables } v_{r1}, v_{r2}, v_{r3} \text{ and add } C_{3-}(v_{r1}, v_{r2}, v_{r3}) \text{ to } F; \\
\hspace{2em} \text{choose 3 distinct variables } v_{r1}, v_{r2}, v_{r3} \text{ and add } C_{3+}(v_{r1}, v_{r2}, v_{r3}) \text{ to } F; \\
\hspace{2em} \text{choose a truth assignment } \hat{t} \in \{0, 1\}^{n} \text{ at random; flip literals whose underlying variables are assigned 0 on } \hat{t}; \\
\text{output } F \text{ and } \hat{t}; \\
\text{end.}
\end{align*}

\begin{figure}[h]
\centering
Fig. 2. Algorithm.
\end{figure}

\textit{u(F, t) is at least }u\text{ for any } t \text{ in } T_{0} (= \{1^{n}\}), T_{1}, T_{n-1} \text{ and } T_{n} (= \{0^{n}\}). \text{ Thus our goal is to show the condition that } X_{t} \text{ is at least } u\text{ for any } t \text{ in } T_{d} \text{ (} 2 \leq d \leq n-2 \text{) with high probability (we remark that } X_{t} \text{ is less than } u(F, t).)

First we consider } \mu_{d} \text{ the expectation of } X_{t} \text{ on arbitrary assignments } t \text{ in } T_{d}.

\textbf{Lemma 1.} For any } d \text{ (} 0 \leq d \leq n)\text{, for any assignment } t \in T_{d}\text{, let } \mu_{d} \text{ be the expected number of clause sets } C_{9-}, C_{9+}, C_{5}, C_{3-} \text{ and } C_{3+} \text{ in } F \text{ such that each set contains at least one clause not satisfied by } t \text{ (the expectation of } X_{t}). \text{ Then }

\begin{align*}
\mu_{d} \geq u - u \frac{d(n-d)}{n(n-1)} + b \frac{d(n-d)}{n-1} + r \frac{d(n-d)}{n(n-1)}.
\end{align*}

By using this lemma, we can use Chernoff bound to bound the probability that } X_{t} \text{ is less than } u \text{ and prove following theorem.}

\textbf{Theorem 3.} For any positive } \varepsilon < 1\text{, the probability that } \hat{t} \text{ is optimal solution of } F \text{ is at least } 1 - \varepsilon \text{ if }

\begin{align*}
r \geq u - bn + 4(n-1) \ln 2n - 2(n-1) \ln \varepsilon + (n-1) \sqrt{2 \ln 2n - \ln \varepsilon \sqrt{2u}}.
\end{align*}

In theoretical analysis, we prove that our instance generator outputs instance with optimal value } u \text{ with high probability if } r \text{ is } u + O(n \ln n) + O(n \sqrt{u \ln n}). \text{ But is such a large number really required? In fact, this upper bound becomes over 1300 even if } u = 5, \text{ } n = 50 \text{ and } \varepsilon = 1. \text{ Thus we compare this theoretical bound with } r \text{ obtained through experiments. Each experiment, we calculate the ratio of instances whose optimal value is } u \text{ for 100 instances.}

First, we estimate whether our algorithm can cover the average optimal value. We fix the number of variables } n \text{ to 50 and } b \text{ to 0. We select } u \text{ from 5, 10, 15 and}
20. Results are in Fig. 3(a). The intervals of results are smaller as opposed to $u$. For each $u$, we pick up two values of $m = 8u + 2r$, the number of total clauses, such that one is the largest $m$ where the ratio is less than 50% and the other is the smallest $m$ where the ratio is larger than 50%. We compare this result with the average optimal value obtained experimentally in Section 2 (Fig. 3(b)). From the figure, we can see that the number of clauses that the ratio of instance with optimal value $u$ is just 50% is near the the number of clauses whose average optimal value is $u$. Thus, we consider that our algorithm generates an instance whose optimal solution is given with high probability if the optimal value is quite smaller than the average optimal value.

Second, we want to estimate whether the main term of upper bound on $r$ is $n \ln n$ or not. We choose $n$ from 40 to 100 at intervals of 10. And we fix $u$ to 5 and $b$ to 0. Results are in Fig. 4(a). Then we also pickup two values of $r$ where the ratio of instances with given optimal solution is just smaller and larger than 0.5 for each $n$. In Fig. 4(b), the vertical axis is value of $r$ where the ratio of instances with given optimal value is just stepping over 0.5, the horizontal axis is the number of variables $n$. From this result, the relation of $r$ and $n$ seems linear or the constant of $n \ln n$ is very small. Hence, we try to fit by linear function, and get $r = 2.6n + 14.8$ (dotted line).
Finally, we change $b$, the number of clause sets $\bigcup_{i=1}^{n} C_5^i$. In the theoretical upper bound, there exists the term $-bn$. This means that adding $b$ sets of $\bigcup_{i=1}^{n} C_5^i$ is equivalent to adding $bn$ random clauses of the each form $C_3^-$ and $C_3^+$ ($2bn$ clauses). We fix the number of variables $n$ to 50 and the optimal value $u$ to 5 again. We choose $b$ from 1, 2 and 3. Results are in Fig. 5, the vertical axis is the ratio of instances whose optimal value is 5 and the horizontal axis is the total number of clauses $m$. We can easily see that $b$ diminishes the total number of clauses. Furthermore, it contribute converging the ratio to 1 much faster.

References

The Euclidean Bottleneck Steiner Tree and Steiner Tree with Minimum Number of Steiner Points

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Abstract. We study variations of Steiner tree problem. Let $P = \{p_1, p_2, \ldots, p_n\}$ be a set of $n$ terminals in the Euclidean plane. For a positive integer $k$, the bottleneck Steiner tree problem (BSTP for short) is to find a Steiner tree with at most $k$ Steiner points such that the length of the longest edges in the tree is minimized. For a positive constant $R$, the Steiner tree problem with minimum number of Steiner points (STP − MSP for short) asks for a Steiner tree such that each edge in the tree has length at most $R$ and the number of Steiner points is minimized. In this paper, we give (1) a ratio-$\sqrt{3} + \epsilon$ approximation algorithm for BSTP, where $\epsilon$ is an arbitrary positive number; (2) a ratio-3 approximation algorithm for STP-MSP with running time $O(n^3)$; (3) a ratio-$\frac{5}{2}$ approximation algorithm for STP-MSP.

1 Introduction

Given a set $P$ of terminals in the Euclidean plane, a Steiner tree on $P$ is an acyclic network interconnecting the terminals. Every vertex in a Steiner tree other than terminals is called a Steiner point.

A variation of Steiner tree problem, named bottleneck Steiner tree problem (BSTP for short) is defined as follows: given a set $P$ of $n$ terminals and a positive integer $k$, we want to find a Steiner tree with at most $k$ Steiner points such that the length of the longest edges in the tree is minimized. Contrary to the classic Steiner tree problem, degree-2 Steiner points are allowed in BSTP. Instead of minimize the total length of the tree, here we want to minimize the length of the longest edges. The problem has applications in the design of wireless communication networks \cite{8}. Due to budget limit, we could put totally $n + k$ stations in the plane, $n$ of them must be located at given points. We would like to have the distances between stations as small as possible. The problem here is
to choose the locations for the other $k$ stations to minimize the longest distance between stations. For other related variations and applications, see [1,2,4,7].

The problem is NP-hard. In [8], it shows that the problem cannot be approximated in polynomial time within performance ratios 2 and $\sqrt{2}$ in the rectilinear plane and the Euclidean plane, respectively. Moreover, a ratio-2 approximation algorithm was given for both the rectilinear plane and the Euclidean plane [8]. For the rectilinear plane, the performance ratio is the best possible. A ratio-1.866 approximation algorithm was given for the Euclidean plane [9]. In this paper, we give a randomized approximation algorithm with performance ratio $\sqrt{3} + \epsilon$ for the Euclidean plane, where $\epsilon$ is an arbitrary positive number.

Another variation of Steiner tree problem considered here is the Steiner tree problem with minimum number of Steiner points, denoted by STP − MSP for short, is defined as follows: Given a set $P$ of $n$ terminals in the Euclidean plane, and a positive constant $R$, STP-MSP asks for a tree $T$ spanning a superset of $P$ such that each edge in the tree has length at most $R$ and the number $C(T)$ of Steiner points is minimized. In the STP-MSP problem, Steiner points of degree 2 are possible. The STP-MSP problem has an important application in wavelength-division multiplexing optical network design [3,5]. In [4], Lin and Xue proved that the STP-MSP problem is NP-hard and gave a ratio-5 approximation algorithm. In [1], Chen, Du et al. showed that the algorithm given by Lin and Xue [4] has performance ratio exactly 4, and they also presented a new $O(n^4)$-time approximation with performance ratio at most 3, where $n$ is the number of given terminals. In this paper, we give a $O(n^3)$-time approximation with performance ratio at most 3, and a randomized approximation with performance ratio at most $\frac{5}{2}$.

A full component of a Steiner tree is a subtree in which each terminal is a leaf and each internal node is a Steiner point. A Steiner tree for $n$ terminals is a $k$-restricted Steiner tree if each full component spans at most $k$ terminals.

## 2 Ratio-$\sqrt{3}$ Approximation Algorithm for BSTP

In this section, we give a ratio-$\sqrt{3}$ approximation algorithm for BSTP in the Euclidean plane.

**Theorem 1** Let $T$ be an optimum Steiner tree for BSTP. Then, there is a 3-restricted Steiner tree with the same number of Steiner points as $T$ such that the longest edge in the tree is at most $\sqrt{3}$ times the optimum.

**Proof:** We assume that $T$ is rooted by arbitrarily selecting a Steiner point as its root. We will modify $T$ bottom up into a 3-restricted Steiner as described in the theorem. Without loss of generality, we assume that $T$ is a full Steiner tree.

We organize the nodes in $T$ level by level (ignoring degree-2 Steiner points). Level 1 is the lowest level. Level $i$ is the level above level $i - 1$. Let $v$ be a node at level 3, $v'$ a child of $v$. If $v'$ is a Steiner point, we can assume that the degree of $v'$ is 3. Otherwise, we can always find two children of $v'$, say $a$ and $b$, such that $\angle av'b \leq 120^\circ$. Without loss of generality, assume $|av'| \geq |bv'|$, and assume
the path $P$ from $a$ to $v'$ has $m$ degree-2 Steiner points. Then we construct a new Steiner tree $T'$ by removing all the edges on $P$, and directly connecting $a$ and $b$ with $m$ degree-2 Steiner points so that the length of each edge in the segment $ab$ is at most $\sqrt{3}$. Now, we need only to consider the tree obtained from $T'$ by removing $a$ and all the degree-2 nodes on the path connecting $a$ and $b$ in $T'$.

In the rest, we assume that the degree of $v'$ is at most $3$. We consider two cases.

**Case 1.** Every edge below $v$ in $T$ has length no more than $1$. We consider the case where $v$ has 4 grandchildren, and leave the case where $v$ has 3-children to the interested readers.

We first consider that $v$ is a degree-3 node. (See Figure 1: (a).) In this case, we can assume that $\angle bv'c > 120^\circ$ and $\angle df e > 120^\circ$. Therefore, we have $\min\{\angle df v, \angle ef v\} < 120^\circ$ and $\min\{\angle cv'v, \angle bv'v\} < 120^\circ$, i.e.,

$$\min\{|dv|, |ev|\} < \sqrt{3} \quad \text{and} \quad \min\{|cv|, |bv|\} < \sqrt{3}. \quad (1)$$

Without loss of generality, assume that

$$\angle df v = \theta = \min\{\angle vv'v, \angle vv'c, \angle df v, \angle ef v\}, \quad \text{and} \quad \angle vv'c \leq \angle vv'b. \quad (2)$$

We will find a point $h$ on edge $vv'$ such that $\max\{|ch|, |bh|, |dh|\} \leq \sqrt{3}$, and construct a new tree by removing nodes $v'$ and $f$, adding edges $ch, cv, dh, bh$, and connecting $d$ and $e$ directly with a Steiner point $w$ on the middle of $de$. (See Figure 1: (b).) Then, we can continue the modification process with $n - 3$ terminals in $P \cup \{v\} - \{b, c, d, e\}$.

By (1) and (2), we know that $|ch| \leq \sqrt{3}$ for any $h$ on the edge $vv'$. So, we need only to choose an $h$ to guarantee that $|bh| \leq \sqrt{3}$ and $|dh| \leq \sqrt{3}$.

First we suppose that $\theta < 90^\circ$, then one can easily check that the point $h$ on edge $vv'$ with $|vh| = 2 - \sqrt{3}$ satisfies the requirement.

Now we suppose that $90^\circ \leq \theta < 120^\circ$. By (2), we know that $\angle bv'v < 360^\circ \ - \ 120^\circ \ - \ \theta = 240^\circ \ - \ \theta$. We choose $h$ to be the point on edge $vv'$ such that $|vh| =
\[ \sqrt{3} - |dv| \text{. By (1), } |vh| > 0. \text{ It is easy to see that} \]
\[ |v'h| \leq 1 - \sqrt{3} + \sqrt{2} - 2 \cos \theta. \quad (3) \]

By triangle inequality,
\[ |dh| \leq |dv| + |vh| = \sqrt{3}. \quad (4) \]

Using (3), we have
\[ |bh|^2 = |bv'|^2 + |v'h|^2 - 2|bv'||v'h| \cos \angle bv'v \]
\[ \leq 1 + (1 - \sqrt{3} + 2 \sin \frac{\theta}{2})^2 + 2(1 - \sqrt{3} + 2 \sin \frac{\theta}{2}) \sin(\theta + 30^\circ) = G(\theta) \]

Set \( H(\theta) = G(\theta) - (\sqrt{3})^2 = -2 + (1 - \sqrt{3} + 2 \sin \frac{\theta}{2})^2 + 2(1 - \sqrt{3} + 2 \sin \frac{\theta}{2}) \sin(\theta + 30^\circ). \)

One can verify directly that \( \frac{d^2H}{d\theta^2}(\theta) > 0 \) for \( 90^\circ \leq \theta \leq 120^\circ \). This together with the fact that \( \frac{d^2H}{d\theta^2}(90^\circ) < 0 \) and \( \frac{d^2H}{d\theta^2}(120^\circ) > 0 \) indicates that \( \frac{d^2H}{d\theta^2}(\theta) \leq \max\{\frac{d^2H}{d\theta^2}(90^\circ), \frac{d^2H}{d\theta^2}(120^\circ)\} < 0 \) for \( \theta \in [90^\circ, 120^\circ] \). So, \( \frac{dH}{d\theta}(\theta) \) is strictly decreasing on \([90^\circ, 120^\circ], \) and then we have for \( \theta \in [90^\circ, 120^\circ], \frac{dH}{d\theta}(\theta) \geq \frac{dH}{d\theta}(120^\circ) > 0 \). This means that \( H(\theta) \) is strictly increasing on \([90^\circ, 120^\circ]. \) So, \( H(\theta) \leq H(120^\circ) = 0 \). Combine this with (4), \( h \) is certainly a correct choice.

**Case 2.** Some edges below \( v \) have length greater than 1. Let \( u \) be a steiner point which is a child of \( v \) and has degree 3, \( x \) and \( y \) the two terminals connected to \( u \).

Without loss of generality, suppose that \(|ux| \leq |uy|\), \( ux \) and \( uy \) have \( l \) and \( k \) Steiner points (both including \( u \), respectively. Let \( z \) be the point on \( uy \) such that \(|uz| = l\). then, we can assume that \( zy \) contains \( k-l \) Steiner points (including \( z \)), and \( ux \) and \( uz \) contain totally \( 2(l-1) \) Steiner points (not including \( u \) and \( z \)).

We directly connect \( x \) and \( z \) and equally insert \([1.155l] - 1 \) Steiner points into \( xz \). Then, the length of each edge on \( xz \) is

\[ \frac{2l}{|1.115l|} \leq \sqrt{3}. \]

After that, we have still \( 2(l-1)-([1.115l])-1 = 2l-1-[1.115l] = [0.885l] - 1 \) Steiner points which can be used to equally break \( ux \) into small edges. Then, by inserting \([0.885l] - 1 \) Steiner points into \( ux \), each edge on \( ux \) has length at most

\[ \frac{l}{[0.885l]} \leq \sqrt{3} \text{ if } l \geq 3. \]

By this operation, \( u \) is changed into a vertex of degree 2 in the new tree, then we can continue the process with \( n - 1 \) terminals in \( P \cup \{u\} \setminus \{x, y\} \).

The case when \( l \leq 2 \) can be proved by using almost the same argument as above. We leave it to the interested readers.
Our algorithm for finding the approximation of an optimal 3-restricted Steiner tree is just the same as that of [9]. It is based on the following theorem about the minimum spanning tree for a weighted 3-hypergraph. A hypergraph \( H = (V, F) \) consists of a set \( V \) (called vertices) and a set \( F \) (called edges) which is an arbitrary family of subsets of \( V \). A weighted hypergraph \( H = (V, F, w) \) is a hypergraph such that each edge \( e \) in \( F \) has a weight \( w(e) \). An \( r \)-hypergraph \( H_r(V, F, w) \) is a weighted hypergraph, each edge has cardinality at most \( r \).

**Theorem 2** [6] There exists a randomized algorithm for the minimum spanning tree problem in 3-hypergraphs running in \( \text{poly}(n, w_{\text{max}}) \) time with probability at least 0.5, where \( n \) is the number of nodes in the hypergraph and \( w_{\text{max}} \) is the largest weight of edges in the hypergraph.

We construct a weighted 3-hypergraph \( H_3(V, F, w) \) from the set \( P \) of terminals. Here \( V = P \), and \( F = \{(a, b)|a \in P \text{ and } b \in P\} \cup \{(a, b, c)|a \in P \text{ and } b \in P \text{ and } c \in P\} \). To obtain the weight of each edge in \( F \), we need to know \( B \), the length of the longest edge in an optimal solution for BSTP. It is hard to find the exact value of \( B \). However, we can find a \( B' \) that is at most \((1 + \epsilon)B\) for any \( \epsilon \) in time \( \text{poly}(n, \epsilon) \) (interested reader can find the detail in [9]).

**Theorem 3** For any given \( \epsilon \), there exists a randomized algorithm that computes a Steiner tree with \( n \) terminals and \( k \) Steiner points such that the longest edge in the tree is at most \( \sqrt{3} + \epsilon \) times of the optimum running in \( \frac{1}{\epsilon} \times \text{poly}(n, k) \) time with probability at least 0.5.

### 3 A Ratio-3 Approximation Algorithm for STP-MSP

In this section, we consider the STP-MSP problem. Given a set \( P \) of \( n \) terminals in the Euclidean plane, and a positive constant \( R \), we want to find a Steiner tree with minimum number of Steiner points such that each edge in the tree has length at most \( R \). In [11], Chen et al. presented an \( O(n^4) \)-time approximation with performance ratio at most 3. With a slightly modification, we may reduce the running time to \( O(n^3) \).

Our algorithm is given in Figure 2.

Since we construct 3-stars in Step 2, the algorithm runs in \( O(n^3) \) time.

For a given set \( P \) of terminals, a minimum spanning tree is a tree interconnecting the terminals in \( P \) with edge between terminals. For a given constant \( R \), a steinerized minimum spanning tree is a tree obtained from a minimum spanning tree by inserting \( \lceil \frac{|ab|}{R} \rceil - 1 \) Steiner points to break each edge \( ab \) into small pieces of length at most \( R \).

Let \( T \) be a Steiner tree and \( e \) be a line segment. \( C(T) \) and \( C(e) \) denote the numbers of Steiner points in \( T \) and \( e \), respectively. \( |e| \) denotes the length of \( e \).

**Lemma 4** [11] Every steinerized minimum spanning tree has the minimum number of Steiner points among steinerized spanning trees.
Algorithm A: 3-approximation Algorithm for STP-MSP

Input A set \( P \) of \( n \) terminals, a positive constant \( R \).
Output A Steiner tree \( T_A \) in which each edge has length at most \( R \).

0. Sort all \( \frac{n(n-1)}{2} \) possible edges between the \( n \) terminals of \( P \) in length increasing order \( e_1, e_2, \ldots, e_{\frac{n(n-1)}{2}} \), and set \( T_A = \emptyset \);
1. for every \( e_i \) such that \( |e_i| \leq R \) do
   if \( e_i \) connects two different connected components of \( T_A \)
   then put \( e_i \) into \( T_A \);
2. for each subset of three terminals \( a, b, c \) respectively in three connected components of \( T_A \) do
   if there exists a point \( s \) within distance \( R \) from \( a, b \) and \( c \)
   then put the 3-star, consisting of three edges \( sa, sb, sc \) into \( T_A \);
3. for every \( e_i \) do
   if \( e_i \) connects two different connected components of \( T_A \)
   then put \( e_i \) into \( T_A \).

Fig. 2. The ratio-3 algorithms.

Lemma 5 \[1\] There exists a shortest optimal Steiner tree \( T^* \) for STP – MSP such that every vertex in \( T^* \) has degree at most five.

Lemma 6 \[1\] Let \( T^* \) be a shortest optimal tree for STP-MSP such that every Steiner point has degree at most five. Let \( T_j \) be a full component of \( T^* \). Then the following hold:

1. The steinerized minimum spanning tree on terminals in \( T_j \) has at most \( 3 \cdot C(T_j) + 1 \) Steiner points.
2. If \( T_j \) contains a Steiner point of degree at most four, then the steinerized minimum spanning tree on terminals in \( T_j \) has at most \( 3 \cdot C(T_j) \) Steiner points.
3. If the steinerized minimum spanning tree on terminals in \( T_j \) has an edge (of length at most \( R \)) between two terminals, then it contains at most \( 3 \cdot C(T_j) \) Steiner points.

From (3) of Lemma 6, we know that if the number of Steiner points contained in a steinerized spanning tree on terminals in a full component \( T_j \) reaches the upper bound \( 3 \cdot C(T_j) + 1 \), then any two terminals are not connected directly by a single edge of length at most \( R \), i.e., there must be a Steiner point between them.

Theorem 7 Let \( T^* \) be an optimal tree for STP-MSP and \( T_A \) an approximation produced by Algorithm A. Then \( C(T_A) \leq 3C(T^*) \).

Proof. Let \( T^S \) be a steinerized minimum spanning tree on all terminals, and let \( k \) be the number of 3-stars produced by Step 2 of Algorithm A. Then

\[ C(T_A) \leq C(T^S) - k. \]
By Lemma 5, we assume that each Steiner point of $T^*$ has degree at most five. Assume that $T^*$ has $h$ full components $T_1, T_2, \ldots, T_h$. For $i = 1, 2$, let $T^{(i)}$ be the components produced by Step $i$ of Algorithm A. We construct a steinerized spanning tree $T$ as follows: Initially, set $T := T^{(1)}$, then for each full component $T_j$ ($1 \leq j \leq h$), add to $T$ the steinerized minimum spanning tree $H_j$ on terminals of $T_j$, if the resulted tree has a cycle, then destroy the cycle by deleting some edges of $H_j$. Without loss of generality, suppose that $T_1, T_2, \ldots, T_g$ ($g \leq h$) are the full components in $T^*$ such that every Steiner point has degree five and $T^{(1)} \cup T_j$ has no cycle. Combining Lemma 4 and Lemma 6 with the fact that for destroying a cycle from $T \cup H_j$, a Steiner point must be removed unless $H_j$ contains an edge between two terminals, we have

$$C(T^S) \leq C(T) \leq 3C(T^*) + g,$$

i.e.,

$$C(T_A) \leq 3C(T^*) + g - k.$$

Suppose that $T^{(1)}$ has $p$ components. Then, $T^{(2)}$ has $p - 2k$ components $C_1, C_2, \ldots, C_{p-2k}$. Now we construct another graph $H$ on all terminals as follows: Initially put all edges of $T^{(1)}$ into $H$, then consider every $T_j$ ($1 \leq j \leq g$). If $T_j$ has a unique Steiner point (this Steiner point connects five terminals which must lie in at most two $C_i$’s), then among the five terminals there are three pairs (edges) of terminals, each pair (edge) lies in the same $C_i$. We add the three edges into $H$. If $T_j$ has at least two Steiner points, then there are two Steiner points each connecting four terminals, and we can also find three pairs (edges) of terminals such that each pair (edge) lies in the same $C_i$. Thus, we can add the three edges into $H$. It is clear that $H$ has at most $p - 3g$ components. Since each components of $H$ is contained by a $C_i$, we have $p - 2k \leq p - 3g$, then $g - k \leq \frac{3g}{2} - k \leq 0$. This ends the proof.

### 4 2.5-Approximation of STP-MSP

In this section, we give a randomized algorithm of ratio-$\frac{5}{2}$ for the STP-MSP problem. Following are some useful Lemmas.

A path $q_1q_2, \ldots, q_m$ in a tree $T$ is called a convex path if for every $i = 1, 2, \ldots, m - 3$, $q_iq_{i+2}$ intersects $q_{i+1}q_{i+3}$. An angle of degree more than 120° is called a big angle. An angle of degree less than or equal to 120° is called a small angle.

**Lemma 8** Let $q_1q_2, \ldots, q_m$ be a convex path. Suppose there are $b$ big angles among $m - 2$ angles $\angle q_1q_2q_3, \angle q_2q_3q_4, \ldots, \angle q_{m-2}q_{m-1}q_m$. Then, $|q_1q_m| \leq (b + 2)R$.

Note that if there is no small angle, $|q_1q_m| \leq (b + 1)R$. Thus, this lemma is useful only when there are many small angles in the convex path.
Lemma 9 In a shortest optimal tree $T$ for $\text{STP} - \text{MSP}$, there are at most two big angles at a point of degree three, there is at most one big angle at a point of degree four, and there is no big angle with degree five.

Let $T^*$ be a shortest optimal tree for $\text{STP} - \text{MSP}$ on $n$ terminals which is a full Steiner tree. Let $s_i$ denote the number of Steiner points of degree $i$ in $T^*$.

Lemma 10 $3s_5 + 2s_4 + s_3 = n - 2$.

Lemma 11 Consider a clockwise tour $F$ of $T^*$ that visits the $n$ terminals in the order $t_1, t_2, \ldots, t_n, t_1$. (See Figure 3.) Then,

(i) the tour $F$ has exactly $n$ convex paths $P_1, P_2, \ldots, P_n$ such that $P_i$ connects two terminals $t_i$ and $t_{i+1}$ ($t_{n+1} = t_1$);
(ii) each angle at a Steiner point appears in those $n$ convex paths exactly once.
(iii) connect the two ends of $P_i$ by an edge $e_i$ and then steinerized $e_i$, $i = 1, 2, \ldots, n$, the total number of Steiner points in any $n-1$ $e_i$'s $\bar{C} = \sum_{i=1}^{n-1} C(e_i)$ is upper bounded as follows:

$$\bar{C} \leq s_4 + 2s_3 + 2s_2 + n - C(e_n) = 3(s_5 + s_4 + s_3) + 2s_2 + 2 - C(e_n)$$

(5)

We denote by $T_F$ the tree consisting of $n$ terminals and $(n-1)$ edges $e_1, e_2, \ldots, e_{n-1}$.

Proof. (i) and (ii) are very easy to see from the structure of $T^*$. Now, we prove (iii). Consider the tour $F$. By Lemma 8 if there are $a_i$ big angles in $P_i$, then there are at most $a_i + 1$ Steiner points on $e_i$, and so the total number of Steiner points in $F$ is at most $n$ plus the number of big angles in $T^*$. By Lemma 9 there are at most $2s_2 + 2s_3 + s_4$ big angles in $T^*$. From Lemma 10 we know that (iii) is valid.

Let $T^*$ be a shortest optimal tree for $\text{STP-MSP}$ which is a full Steiner tree on $n$ terminals. Without loss of generality, we assume that $T^*$ has Steiner points

Fig. 3. The tour $F$ that visits all the terminals
of degree at least three. Selecting an arbitrary Steiner point of degree at least three as the root of $T^*$, we get a rooted tree. A **good point** $t$ in $T^*$ is a Steiner point that is adjacent to some terminals and satisfies one of the following:

(i) (type (1)) $t$ has three or more terminals as children;
(ii) (type (2)) $t$ has two terminals as children and the degree of $t$ is 4;
(iii) (type (3)) $t$ is a point of degree 3.

Note that a good point is of degree at least 3. A **bad point** is a Steiner point of degree at least 3 in $T^*$ that is not a good point.

**Theorem 12** There is a 3-restricted Steiner tree such that each edge has length at most $R$ and the Steiner points is at most $\frac{5}{2}$ times the optimum.

Now, we focus on the computation of an optimal 3-restricted tree.

Let $H_3(V,F,W)$ be a weighted 3-hypergraph, where $V = P$, $F = \{(a,b)|a \in V \text{ and } b \in V\} \cup \{(a,b,c)|a \in V, b \in V \text{ and } c \in V\}$, and for each edge $e \in F$, $w(e)$ is the smallest number of Steiner points to form an optimal solution of the STP-MSP problem on the terminals in $e$.

Given three points $a, b$ and $c$ on the Euclidean plane, let $s$ be the Steiner point which minimizes $(|sa| + |sb| + |sc|)$, and let $k$ be the number of Steiner points in an optimum solution $T$ of $STP-MSP$ on $\{a, b, c\}$ with constant $R$.

**Lemma 13**

\[
\left\lfloor \frac{|sa|}{R} \right\rfloor + \left\lfloor \frac{|sb|}{R} \right\rfloor + \left\lfloor \frac{|sc|}{R} \right\rfloor - 2 \geq k \geq \left\lfloor \frac{|sa|}{R} \right\rfloor + \left\lfloor \frac{|sb|}{R} \right\rfloor + \left\lfloor \frac{|sc|}{R} \right\rfloor - 2. \tag{6}
\]

**Proof.** By steinerizing the optimum Steiner tree, we get a solution of $STP-MSP$ on $\{a, b, c\}$ with exactly $\left\lfloor \frac{|sa|}{R} \right\rfloor + \left\lfloor \frac{|sb|}{R} \right\rfloor + \left\lfloor \frac{|sc|}{R} \right\rfloor - 2$ Steiner points.

Let $|T|$ be the total length of $T$, which is the sum of the length of edges of $T$. Then

\[
(k - 1) \cdot R + 3R \geq |T| \geq |sa| + |sb| + |sc|;
\]

i.e.,

\[
k + 2 \geq \frac{|sa| + |sb| + |sc|}{R} \geq \left\lfloor \frac{|sa| + |sb| + |sc|}{R} \right\rfloor \geq \left\lfloor \frac{|sa|}{R} \right\rfloor + \left\lfloor \frac{|sb|}{R} \right\rfloor + \left\lfloor \frac{|sc|}{R} \right\rfloor.
\]

Therefore, (6) holds.

**Lemma 13** gives an upper bound on the cost of $(a, b, c)$.

**Lemma 14** Testing whether three circles has a point in common can be done in constant time.

For any given points $a, b$ and $c$ on the Euclidean plane, one can find the minimum Steiner tree on $\{a, b, c\}$ in constant time. Let $q_{a,b,c}$ be the number of Steiner points used to steinerize the optimum Steiner tree on $\{a, b, c\}$, and $q_P = \max\{q_{a,b,c}|\{a, b, c\} \subset P\}$. Then, by Lemma 13 and Lemma 14, the weight $W$ of $H_3(V,F,W)$ can be calculated in $O(n^3q_P^2)$ time. By Theorem 2 and Theorem 12 we have
$\frac{5}{2}$-approximation Algorithm for STP-MSP

Input A set $P$ of $n$ terminals in the Euclidean plane, a positive constant $R$
Output A 3-restricted Steiner tree $T$ in which each edge has length at most $R$.
1. Construct a weighted hypergraph $H_3(V, F, w)$.
2. Call the randomized algorithm in [6] to compute a minimum spanning tree $T'$ for $H_3(V, F, w)$;
3. Replace every edge $f$ of the MST $T'$ of $H_3(V, F, w)$ with a Steiner tree with $w(f)$ Steiner points such that the maximum length of each edge in the tree is at most $R$ and output the obtained tree.

Fig. 4. The complete algorithm.

Theorem 15 Given a set $P$ of $n$ terminals and a positive constant $R$, there exists a randomized algorithm that computes a solution of $STP - MSP$ on $P$ such that the number of Steiner points is at most $\frac{5}{2}$ times of the optimum running in poly$(n, q_P)$ time with probability at least 0.5.

The complete algorithm is given in Figure 4.

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References
An FPTAS for Weight-Constrained Steiner Trees in Series-Parallel Graphs

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Abstract. In this paper, we study the problem of computing a minimum cost Steiner tree subject to weight constraint in a series-parallel graph where each edge has a nonnegative integer cost and a nonnegative integer weight. We present a fully polynomial time approximation scheme for this NP-complete problem.

1 Introduction

A computer network is often modeled by an undirected graph $G = (V,E)$ where $V$ is the set of vertices and $E$ is the set of edges. The traditional Steiner tree problem associates an edge cost $c(e) \geq 0$ with each edge $e \in E$ and asks for a minimum cost subgraph of $G$ spanning a given subset $T \subseteq V$ of target vertices. Such problems find important applications in computer networks and have been studied by many researchers \cite{1,4,10,13,15}. We refer the readers to \cite{8,11} for details.

In this traditional network model, there is only one cost corresponding to each edge. In real-life problems, each edge may have a weight besides a cost, and we want to find a minimum cost subgraph of $G$ spanning the vertices in $T$ with a total weight no more than a given weight constraint $W$. We call this problem the weight constrained minimum cost Steiner tree problem (WCSTP).

Formally, we generalize the traditional network model to allow two independent edge weighting functions: with each edge $e \in E$, there is an associated integer cost $c(e) \geq 0$ and an associated integer weight $w(e) \geq 0$. Let $G'$ be a subgraph of $G$. The cost of $G'$, denoted by $c(G')$, is the sum of the edge costs of $G'$. The weight of $G'$, denoted by $w(G')$, is the sum of the edge weights of $G'$. Given a set of target vertices $T \subseteq V$ and an integer weight constraint $W \geq 0$, we are interested in computing a minimum cost tree subgraph $T$ of $G$ which spans the vertices in $T$ subject to the constraint that $w(T) \leq W$. We call such a tree a weight constrained minimum cost Steiner minimum tree.

When $w(e) = 0$ for every edge $e \in E$, the WCSTP becomes the well known Steiner tree problem. Since the Steiner tree problem is NP-hard in the strong sense \cite{6}, the WCSTP is also NP-hard in strong sense. In this paper, we are
interested in WCSTP on a very important class of sparse networks known as series-parallel graphs (S-P graphs), which are subgraphs of 2-trees defined below.

Following [13], a 2-tree can be defined recursively as follows, and all 2-trees may be obtained in this way. A triangle (a complete graph on three vertices) is a 2-tree. Given a 2-tree and an edge \{x, y\} of the 2-tree, we can add a new vertex z adjacent to both x and y; the result is a 2-tree. A series-parallel graph (also known as a partial 2-tree) is a subgraph of a 2-tree. With this definition, one can see that a 2-tree on n vertices has 2n − 3 edges and n − 2 triangles.

Farley [5] demonstrated that 2-trees are isolated failure immune (IFI) networks. Wald and Colbourn [13] showed that a minimum IFI network is a 2-tree. This fact made 2-trees an important class of fault tolerant networks. Wald and Colbourn [13] also showed that the Steiner tree problem on a series-parallel graph is tractable, presenting a linear time algorithm. In a recent paper, Colbourn and Xue [3] showed that the grade of service Steiner tree problem [16] is also tractable on a series-parallel graph.

In this paper, we show that the weight constrained minimum cost Steiner tree problem is NP-hard even when the graph \( G \) is a series-parallel graph. We then present a fully polynomial time approximation scheme for computing a weight constrained minimum cost Steiner tree in a series-parallel graph.

2 WCSTP in S-P Graphs Is Intractable

When the number of target vertices is zero or one, the WCSTP is trivial. When there are only two vertices in \( T \), the WCSTP becomes the well-known weight constrained shortest path problem (WCSPP) which is known to be NP-hard [6]. To the best of our knowledge, no one has ever addressed the complexity of WCSTP or WCSPP in an S-P graph specifically. In a recent paper [14], Wang and Crowcroft presented an NP-hardness proof for the WCSPP problem in general graphs. We point out that the NP-hardness proof of Wang and Crowcroft, although targeted at the problem in general graphs, also proves that the WCSPP is NP-hard in an S-P graph. Since the WCSPP in an S-P graph is NP-hard, the more general WCSTP in an S-P graph is also NP-hard. Therefore we have the following hardness result (for the completeness of the current paper, we include an NP-hardness proof of the WCSPP in an S-P graph in the appendix).

**Theorem 1.** The WCSTP in an S-P graph is NP-hard.

Note that in the definition of the WCSTP, the weight and cost of an edge are symmetric. Therefore we may also talk about cost constrained minimum weight Steiner tree problem (CCSTP) which asks for a minimum weight tree subject to cost constraint. Given the hardness of the problem, we are interested in designing efficient approximation algorithms for this problem. In the next section, we will present a pseudo-polynomial time algorithm for computing a cost constrained minimum weight Steiner tree in an S-P graph. We will then apply standard techniques of scaling and rounding to turn the pseudo-polynomial time algorithm into a fully polynomial time approximation scheme (FPTAS) for weight constrained minimum cost Steiner trees.
3 A Pseudo Polynomial Time Algorithm for CCSTP

In this section, we will present a pseudo polynomial time algorithm for computing a cost constrained minimum weight Steiner tree. Given the set \( T \subseteq V \) and a nonnegative integer \( \zeta \), the algorithm computes a minimum weight Steiner tree among those whose cost is bounded by \( \zeta \), in \( O(n\zeta^2) \) time, where \( n \) is the number of vertices in the series-parallel graph.

Our algorithm for computing a weight-constrained Steiner tree operates in two phases. The first completes the graph to a 2-tree, as was done in [13], where each added edge has a weight of zero and a cost that is larger than the cost constraint \( \zeta \). This ensures that the added edges will never be chosen in a cost constrained minimum weight Steiner tree.

The second phase involves finding cost constrained minimum weight Steiner trees in 2-trees.

We are given a 2-tree \( G \) with target set \( T \subseteq V \). With each (directed) arc \( \alpha = (x,y) \) corresponding to an (undirected) edge \( \{x,y\} \) of \( G \), we will associate \( 6\zeta + 6 \) measures, which summarize the weight incurred so far in the subgraph \( S \) which has been reduced onto the edge \( \{x,y\} \):

- \( st(\alpha, \xi) \) is the minimum weight of a Steiner tree for \( S \), with a cost of at most \( \xi \), in which \( x \) and \( y \) appear in the same tree, \( \xi = 0,1,2,\ldots,\zeta \).
- \( dt(\alpha, \xi) \) is the minimum weight of two disjoint trees for \( S \) including all targets, one tree involving \( x \) and the other involving \( y \), whose total cost is no more than \( \xi \), \( \xi = 0,1,2,\ldots,\zeta \).
- \( yn(\alpha, \xi) \) is the minimum weight of a Steiner tree for \( S \), which includes \( x \) but not \( y \), whose cost is at most \( \xi \), \( \xi = 0,1,2,\ldots,\zeta \).
- \( ny(\alpha, \xi) \) is the minimum weight of a Steiner tree for \( S \), which includes \( y \) but not \( x \), whose cost is at most \( \xi \), \( \xi = 0,1,2,\ldots,\zeta \).
- \( nn(\alpha, \xi) \) is the minimum weight of a Steiner tree for \( S \), which includes neither \( x \) nor \( y \), whose cost is at most \( \xi \), \( \xi = 0,1,2,\ldots,\zeta \).
- \( none(\alpha, \xi) \) is the weight of omitting all vertices of \( S \) from a Steiner tree whose cost is at most \( \xi \), \( \xi = 0,1,2,\ldots,\zeta \) (note that there is a big penalty here if \( S \) contains any target vertices).

Let \( \beta = (x,y) \) be the other arc corresponding to the edge \( \{x,y\} \). Then the \( 6\zeta + 6 \) measures associated with \( \beta \) relates to the measures associated with \( \alpha \) in the following way:

\[
\begin{align*}
st(\beta, \xi) &= st(\alpha, \xi)
dt(\beta, \xi) &= dt(\alpha, \xi)
yn(\beta, \xi) &= yn(\alpha, \xi)
nn(\beta, \xi) &= nn(\alpha, \xi)
one(\beta, \xi) &= none(\alpha, \xi)
\end{align*}
\]

Let BIG be an integer which is bigger than the sum of all the edge weights in the given graph. Initially, no reduction of the graph has been done. We set the initial measures for each arc \( \alpha = (x,y) \) corresponding to an edge \( e = \{x,y\} \) as follows, for \( \xi = 0,1,2,\ldots,\zeta \).

Our algorithm (call it Algorithm 1) will use these initial arc measures to recompute the arc measures as the graph is reduced. The graph is reduced by
The algorithm stops when there is no degree-2 vertex left. Suppose at some point in the algorithm there is a triangle \( \{x, y, z\} \) in which \( y \) is a degree-2 vertex and that costs have been computed for both \( (x, y) \) and \( (y, z) \). Let \( L = (x, y) \), \( R = (y, z) \), \( M = (x, z) \). We will update the measures at \( M \) using the following rules:

- For \( \xi = 0, 1, 2, \ldots, \zeta \), let \( w_{\xi, st} \) be the minimum of the following four sets of numbers:
  \[
  \{\text{st}(M, \xi M) + \text{st}(L, \xi L) + \text{dt}(R, \xi R) | \xi L + \xi R + \xi M = \xi, \xi L \geq 0, \xi R \geq 0, \xi M \geq 0\};
  \]
  \[
  \{\text{st}(M, \xi M) + \text{dt}(L, \xi L) + \text{st}(R, \xi R) | \xi L + \xi R + \xi M = \xi, \xi L \geq 0, \xi R \geq 0, \xi M \geq 0\};
  \]
  \[
  \{\text{st}(M, \xi M) + \text{yn}(L, \xi L) + \text{nyn}(R, \xi R) | \xi L + \xi R + \xi M = \xi, \xi L \geq 0, \xi R \geq 0, \xi M \geq 0\};
  \]
  \[
  \{\text{dt}(M, \xi M) + \text{st}(L, \xi L) + \text{st}(R, \xi R) | \xi L + \xi R + \xi M = \xi, \xi L \geq 0, \xi R \geq 0, \xi M \geq 0\}.
  \]

- For \( \xi = 0, 1, 2, \ldots, \zeta \), let \( w_{\xi, dt} \) be the minimum of the following three sets of numbers:
  \[
  \{\text{dt}(M, \xi M) + \text{st}(L, \xi L) + \text{dt}(R, \xi R) | \xi L + \xi R + \xi M = \xi, \xi L \geq 0, \xi R \geq 0, \xi M \geq 0\};
  \]
  \[
  \{\text{dt}(M, \xi M) + \text{dt}(L, \xi L) + \text{st}(R, \xi R) | \xi L + \xi R + \xi M = \xi, \xi L \geq 0, \xi R \geq 0, \xi M \geq 0\};
  \]
  \[
  \{\text{dt}(M, \xi M) + \text{yn}(L, \xi L) + \text{nyn}(R, \xi R) | \xi L + \xi R + \xi M = \xi, \xi L \geq 0, \xi R \geq 0, \xi M \geq 0\}.
  \]

- For \( \xi = 0, 1, 2, \ldots, \zeta \), let \( w_{\xi, yn} \) be the minimum of the following two sets of numbers:
  \[
  \{\text{yn}(M, \xi M) + \text{yn}(L, \xi L) + \text{nyn}(R, \xi R) | \xi L + \xi R + \xi M = \xi, \xi L \geq 0, \xi R \geq 0, \xi M \geq 0\};
  \]
  \[
  \{\text{yn}(M, \xi M) + \text{st}(L, \xi L) + \text{yn}(R, \xi R) | \xi L + \xi R + \xi M = \xi, \xi L \geq 0, \xi R \geq 0, \xi M \geq 0\}.
  \]

- For \( \xi = 0, 1, 2, \ldots, \zeta \), let \( w_{\xi, ny} \) be the minimum of the following two sets of numbers:
  \[
  \{\text{ny}(M, \xi M) + \text{ny}(L, \xi L) + \text{st}(R, \xi R) | \xi L + \xi R + \xi M = \xi, \xi L \geq 0, \xi R \geq 0, \xi M \geq 0\};
  \]
  \[
  \{\text{ny}(M, \xi M) + \text{none}(L, \xi L) + \text{yn}(R, \xi R) | \xi L + \xi R + \xi M = \xi, \xi L \geq 0, \xi R \geq 0, \xi M \geq 0\}.
  \]

- For \( \xi = 0, 1, 2, \ldots, \zeta \), let \( w_{\xi, nn} \) be the minimum of the following four sets of numbers:
  \[
  \{\text{nn}(M, \xi M) + \text{none}(L, \xi L) + \text{none}(R, \xi R) | \xi L + \xi R + \xi M = \xi, \xi L \geq 0, \xi R \geq 0, \xi M \geq 0\}.
  \]
\{\text{none}(M,\xi_M) + \text{nn}(L,\xi_L) + \text{none}(R,\xi_R)|\xi_L + \xi_R + \xi_M = \xi, \xi_L \geq 0, \xi_R \geq 0, \xi_M \geq 0\};
\{\text{none}(M,\xi_M) + \text{none}(L,\xi_L) + \text{nn}(R,\xi_R)|\xi_L + \xi_R + \xi_M = \xi, \xi_L \geq 0, \xi_R \geq 0, \xi_M \geq 0\};
\{\text{none}(M,\xi_M) + \text{nn}(L,\xi_L) + \text{nn}(R,\xi_R)|\xi_L + \xi_R + \xi_M = \xi, \xi_L \geq 0, \xi_R \geq 0, \xi_M \geq 0\}.

- For $\xi = 0, 1, 2, \ldots, \zeta$, let $w_{\xi,\text{none}}$ be the minimum of the following set of numbers:
\{\text{none}(M,\xi_M) + \text{none}(L,\xi_L) + \text{none}(R,\xi_R)|\xi_L + \xi_R + \xi_M = \xi, \xi_L \geq 0, \xi_R \geq 0, \xi_M \geq 0\}.

- Delete the vertex $y$ from the graph. For $\xi = 0, 1, 2, \ldots, \zeta$, update the values of $\text{st}(M,\xi)$, $\text{dt}(M,\xi)$, $\text{yn}(M,\xi)$, $\text{ny}(M,\xi)$, $\text{nn}(M,\xi)$, and $\text{none}(M,\xi)$ to $w_{\xi,\text{st}}$, $w_{\xi,\text{dt}}$, $w_{\xi,\text{yn}}$, $w_{\xi,\text{ny}}$, $w_{\xi,\text{nn}}$, and $w_{\xi,\text{none}}$, respectively.

**Theorem 2.** When the 2-tree is reduced to a single edge $\alpha = \{x, y\}$, the minimum of the four values $\text{st}(\alpha, \zeta)$, $\text{yn}(\alpha, \zeta)$, $\text{ny}(\alpha, \zeta)$, $\text{nn}(\alpha, \zeta)$ is the minimum weight of a Steiner tree interconnecting all the vertices in $T$ with a cost not more than $\zeta$, where a weight of BIG or more indicates the nonexistence of a cost constrained Steiner tree.

**Proof.** We note that the measures of an edge $\alpha = \{x, y\}$ are initialized correctly, and are updated correctly after each reduction. For example, for any $\xi \in \{0, 1, 2, \ldots, \zeta\}$ and an edge $\alpha = \{x, y\}$, $\text{st}(\alpha, \xi)$ is the minimum weight of a Steiner tree for $S$ whose cost is no more than $\xi$, in which $x$ and $y$ appear in the same tree.

When the graph is reduced to a single arc $\alpha = \{x, y\}$, $\text{st}(\alpha, \zeta)$ is the minimum weight of a cost constrained Steiner tree interconnecting $x$, $y$, and all the vertices in $T$; $\text{yn}(\alpha, \zeta)$ is the minimum weight of a cost constrained Steiner tree interconnecting $x$ (but not $y$) and all the vertices in $T - \{y\}$; $\text{ny}(\alpha, \zeta)$ is the minimum weight of a cost constrained Steiner tree interconnecting $y$ (but not $x$) and all the vertices in $T - \{x\}$; $\text{nn}(\alpha, \zeta)$ is the minimum weight of a cost constrained Steiner tree interconnecting $y$ (but not $x$) and all the vertices in $T - \{x, y\}$. Note that leaving out a target vertex from the tree receives a penalty of at least BIG. Therefore the minimum of the above four values corresponds to the minimum weight of a cost constrained Steiner tree. \qed

**Theorem 3.** Algorithm 1 can be implemented in $O(n\zeta^2)$ time, where $\zeta$ is the cost constraint and $n$ is the number of vertices in the graph.

**Proof.** We will show that the update of measurements can be accomplished in $O(\zeta^2)$ time for each reduction. Since there are $O(n)$ reductions, this would complete the proof of the theorem.

Since there are $O(\zeta)$ choices for each of $\xi$, $\xi_M$ and $\xi_L$ ($\xi_R = \xi - \xi_M - \xi_L$), a simple-minded implementation would require $O(\zeta^3)$ time to update the measures for each reduction.
Let us denote $\xi_L + \xi_R$ by $\xi_{LR}$. Although there are $O(\zeta^2)$ different choices for $(\xi_L, \xi_R)$, there are only $O(\zeta)$ different choices for $\xi_{LR}$. Using $O(\zeta^2)$ time, we can compute the following $12(\zeta + 1)$ measures for all $\xi_{LR} \in \{0, 1, 2, \ldots, \zeta\}$.

$$LR(\xi_{LR}, st, st) = \min\{st(L, \xi_L) + st(R, \xi_{LR} - \xi_L) | 0 \leq \xi_L \leq \xi_{LR}\};$$

$$LR(\xi_{LR}, st, dt) = \min\{st(L, \xi_L) + dt(R, \xi_{LR} - \xi_L) | 0 \leq \xi_L \leq \xi_{LR}\};$$

$$LR(\xi_{LR}, dt, st) = \min\{dt(L, \xi_L) + st(R, \xi_{LR} - \xi_L) | 0 \leq \xi_L \leq \xi_{LR}\};$$

$$LR(\xi_{LR}, yn, ny) = \min\{yn(L, \xi_L) + ny(R, \xi_{LR} - \xi_L) | 0 \leq \xi_L \leq \xi_{LR}\};$$

$$LR(\xi_{LR}, yn, none) = \min\{yn(L, \xi_L) + none(R, \xi_{LR} - \xi_L) | 0 \leq \xi_L \leq \xi_{LR}\};$$

$$LR(\xi_{LR}, st, yn) = \min\{st(L, \xi_L) + yn(R, \xi_{LR} - \xi_L) | 0 \leq \xi_L \leq \xi_{LR}\};$$

$$LR(\xi_{LR}, ny, st) = \min\{ny(L, \xi_L) + st(R, \xi_{LR} - \xi_L) | 0 \leq \xi_L \leq \xi_{LR}\};$$

$$LR(\xi_{LR}, ny, none) = \min\{none(L, \xi_L) + ny(R, \xi_{LR} - \xi_L) | 0 \leq \xi_L \leq \xi_{LR}\};$$

$$LR(\xi_{LR}, none, none) = \min\{none(L, \xi_L) + none(R, \xi_{LR} - \xi_L) | 0 \leq \xi_L \leq \xi_{LR}\};$$

$$LR(\xi_{LR}, nn, none) = \min\{nn(L, \xi_L) + none(R, \xi_{LR} - \xi_L) | 0 \leq \xi_L \leq \xi_{LR}\};$$

$$LR(\xi_{LR}, none, nn) = \min\{none(L, \xi_L) + nn(R, \xi_{LR} - \xi_L) | 0 \leq \xi_L \leq \xi_{LR}\};$$

$$LR(\xi_{LR}, ny, yn) = \min\{ny(L, \xi_L) + yn(R, \xi_{LR} - \xi_L) | 0 \leq \xi_L \leq \xi_{LR}\}.$$  

After those $12(\zeta + 1)$ measures are computed, only $O(\zeta^2)$ time are necessary to compute the updated values for $st(M, \xi)$, $dt(M, \xi)$, $yn(M, \xi)$, $ny(M, \xi)$, $nn(M, \xi)$, $none(M, \xi)$, for $\xi = 0, 1, 2, \ldots, \zeta$. For example, $w_{\xi, yn}$ can be computed as the minimum of

$$\min\{yn(M, \xi_M) + LR(\xi - \xi_M, yn, none) | 0 \leq \xi_M \leq \xi\}$$

and

$$\min\{yn(M, \xi_M) + LR(\xi - \xi_M, st, yn) | 0 \leq \xi_M \leq \xi\}.$$  

This shows that Algorithm 1 can be implemented in $O(n\zeta^2)$ time. \hfill \square

We point out that the weight constrained Steiner tree can be constructed in $O(n\zeta)$ extra time time if we perform some bookkeeping operations (recording how the minimum was achieved) during the reductions.

### 4 An FPTAS for WCSTP

We use standard techniques of scaling and rounding \cite{27, 9, 12, 17} to turn the pseudo polynomial time algorithm for CCSTP into a fully polynomial time approximation scheme for WCSTP.

Let us use $c(\mathcal{T}, \mathcal{W})$ to denote the minimum cost of a Steiner tree spanning the targets in $\mathcal{T}$ with a weight of no more than $\mathcal{W}$. Given a positive real number $\mathcal{C}$, deciding whether $c(\mathcal{T}, \mathcal{W}) > \mathcal{C}$ or $c(\mathcal{T}, \mathcal{W}) \leq \mathcal{C}$ is NP-hard. Using the standard technique of scaling and rounding \cite{27, 9, 12, 17}, we can decide, in fully polynomial time, whether $c(\mathcal{T}, \mathcal{W}) > \mathcal{C}$ or $c(\mathcal{T}, \mathcal{W}) < (1 + \epsilon)\mathcal{C}$, for any given constant $\epsilon > 0$. This technique will play an important role in our FPTAS for computing
a weight constrained minimum cost Steiner tree in an S-P graph. We describe
this approximate testing in Algorithm 2 as TEST.

Algorithm 2 TEST($\mathcal{C}, \epsilon$)

\begin{itemize}
  \item \textbf{Step} 1. Set $\theta := \frac{2^{\frac{n}{\epsilon}} - 3}{\epsilon \times \mathcal{C}}$; Let $c_\theta$ be the scaled edge cost function such that $c_\theta(e) = [c(e) \times \theta]$ for $e \in E$; Set $\zeta := \mathcal{C} \times \theta$;
  \item \textbf{Step} 2. Apply Algorithm 1 using the scaled edge cost function $c_\theta$ instead of the original edge cost function $c$;
    \begin{itemize}
      \item if the weight of the cost constrained Steiner tree is no more than $\mathcal{W}$ then
        \begin{itemize}
          \item output YES;
        \end{itemize}
      \item else
        \begin{itemize}
          \item output NO;
        \end{itemize}
    \end{itemize}
\end{itemize}

\textbf{Theorem 4.} Let us be given the target set $\mathcal{T}$, the weight constraint $\mathcal{W}$, the positive real numbers $\mathcal{C}$ and $\epsilon$. If TEST($\mathcal{C}, \epsilon$) = NO, then $c(\mathcal{T}, \mathcal{W}) > \mathcal{C}$. If TEST($\mathcal{C}, \epsilon$) = YES, then $c(\mathcal{T}, \mathcal{W}) < (1 + \epsilon) \times \mathcal{C}$. In addition, the worst-case time complexity of TEST($\mathcal{C}, \epsilon$) is $O\left(\frac{n^3}{\epsilon^2}\right)$.

\textbf{Proof.} Let $T$ be a tree subgraph in $G$. Let $c(T) = \sum_{e \in T} c(e)$ and $c_\theta(T) = \sum_{e \in T} c_\theta(e)$. Since the number of edges in $T$ is at most $2n - 3$, we can prove that $c_\theta(T) \leq \zeta$ implies $c(T) \leq \mathcal{C}(1 + \epsilon)$ and that $c_\theta(T) > \zeta$ implies $c(T) > \mathcal{C}$, where $\theta$ and $\zeta$ are as defined in Algorithm 2.

Assume that TEST($\mathcal{C}, \epsilon$) = NO. Then we know that for any tree $T$ spanning the targets in $\mathcal{T}$ with a weight no more than $\mathcal{W}$, we must have $c_\theta(T) > \zeta$, which in turn implies $c(T) > \mathcal{C}$. This says that $c(\mathcal{T}, \mathcal{W}) > \mathcal{C}$.

Now assume that TEST($\mathcal{C}, \epsilon$) = YES. Then we know that there is a tree $T$ spanning the targets in $\mathcal{T}$ such that $w(T) \leq \mathcal{W}$ and $c_\theta(T) \leq \zeta$. Note that $c_\theta(T) \leq \zeta$ implies $c(T) < \mathcal{C}(1 + \epsilon)$. This says that $c(\mathcal{T}, \mathcal{W}) > \mathcal{C}(1 + \epsilon)$.

The time complexity of TEST follows directly from Theorem 3. \hfill $\Box$

We will use LB and UB to denote lower and upper bounds on $c(\mathcal{T}, \mathcal{W})$. Our FPTAS starts with efficiently computable values of LB and UB and uses bisection to reduce the ratio $\frac{UB}{LB}$.

An initial value of LB can be computed as follows. Use Algorithm 1 to compute a minimum weight Steiner tree $T$ whose cost is 0. This clearly can be done in polynomial time. If the weight of this tree is no more than $\mathcal{W}$, it is also a weight constrained minimum cost Steiner tree and we are done. Otherwise, $c(\mathcal{T}, \mathcal{W}) \geq 1$ and we can use 1 as the initial value of LB.

An initial value of UB can also be computed easily. Use the algorithm in LB to compute a minimum weight (ignore the cost) Steiner tree $T$ spanning the target vertices in $\mathcal{T}$. If $w(T) > \mathcal{W}$, there is no Steiner tree whose weight is bounded by $\mathcal{W}$. When $w(T) \leq \mathcal{W}$, we may use $c(T)$ as the initial value of UB.

Let B be some chosen real number which is grater than $1 + \epsilon/3$. We will apply bisection to drive the ratio $UB/LB$ down to some number below B. Suppose that our lower bound LB and upper bound UB are such that $UB > LB \times B > LB \times (1 + \epsilon/3)$. Let $\mathcal{C} = \sqrt{\frac{UB \times LB}{1 + \epsilon/3}}$. If TEST($\mathcal{C}, \epsilon/3$) = NO, then $\mathcal{C}$ is also a lower bound for $c(\mathcal{T})$. 

If $\text{TEST}(\mathcal{C}, \epsilon/3) = \text{YES}$, then $(1 + \epsilon/3) \times \mathcal{C}$ is also an upper bound for $c(T)$. Therefore, the ratio of the new upper bound over the new lower bound will be no more than $\sqrt{\frac{\text{UB}}{\text{LB}}} \times (1 + \epsilon/3)$. Let us call the above process an iteration. Note that such an iteration can be accomplished in fully polynomial time (by Theorem 4). Furthermore, the ratio of the upper-bound over the lower bound can be reduced to a number below $B$ in polynomial number of iterations (polynomial in the input size of the given instance and $\frac{1}{\epsilon}$). This analysis leads to our FPTAS which is given in Algorithm 3.

**Algorithm 3**

**FPTAS for weight constrained minimum cost Steiner tree in an S-P graph.**

<table>
<thead>
<tr>
<th>Step</th>
<th>Description</th>
</tr>
</thead>
<tbody>
<tr>
<td>1</td>
<td>set $\mathcal{B} := (1 + \epsilon/3) \times (1 + \epsilon/3)$; set LB and UB to their initial values;</td>
</tr>
<tr>
<td>2</td>
<td>if UB $\leq \mathcal{B} \times$ LB then goto Step 3; else let $\mathcal{C} := \sqrt{\frac{\text{UB} \times \text{LB}}{1 + \epsilon/3}}$; if $\text{TEST}(\mathcal{C}, \epsilon/3) = \text{NO}$, set LB $= \mathcal{C}$; if $\text{TEST}(\mathcal{C}, \epsilon/3) = \text{YES}$, set UB $= \mathcal{C} \times (1 + \epsilon/3)$; goto Step 2;</td>
</tr>
<tr>
<td>3</td>
<td>set $\theta := \frac{2n - 3}{\text{LB} \times \epsilon/3}$; $\zeta := \theta \times$ UB; set $c_\theta(e) := \lfloor \theta \times c(e) \rfloor$ for every $e \in E$; apply Algorithm 1 to compute a cost constrained minimum weight Steiner tree using the scaled cost function $c_\theta$.</td>
</tr>
</tbody>
</table>

**Theorem 5.** If there is no weight constrained Steiner tree spanning the target vertices in $T$, we will find this out during our computation of the initial value of UB. If $T$ has a weight constrained Steiner tree, Algorithm 3 finds a weight constrained Steiner tree $T$ such that $w(T) \leq W$ and $c(T) \leq (1 + \epsilon) \times c(T, W)$. Furthermore, the time complexity of Algorithm 3 is $O\left(\frac{n^3}{\epsilon^2} \times (\log S + \log(\frac{1}{\epsilon}))\right)$, where $S$ is the input size of the given instance.

**Proof.** The claim follows from standard techniques [2, 7, 17] and Theorems 2-4. \qed

## 5 Conclusions

In this paper, we have studied the weight constrained minimum cost Steiner tree problem on a very important class of sparse graphs–series-parallel graphs. Although most Steiner tree problems are tractable on series-parallel graphs, it is shown that the weight constrained minimum cost Steiner tree problem on series-parallel graphs is NP-hard. On the positive side, we have presented a fully polynomial time approximation scheme for this problem, which has applications to computer networks as well as computational biology. A more challenging problem is the weight constrained minimum cost Steiner tree problem on a general graph.
Acknowledgments

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17. G.L. Xue and W. Xiao, An FPTAS for minimum cost delay constrained multicast tree under a Steiner topology, with applications, submitted for publication.
A NP-Completeness of WCSTP in S-P Graphs

In this section, we prove that the weight constrained shortest path problem \([7]\) in S-P graphs is NP-hard. After we have arrived at this proof, we found that the proof by Wang and Crowcroft \([14]\) also proves the same fact. This proof is included here as an aid to the referees.

**Theorem 6.** The weight constrained shortest path problem in series-parallel graphs is NP-hard.

**Proof.** Now we transform PARTITION to the weight constrained shortest path problem in S-P graphs.

Suppose we have a set \(S\) with \(n\) elements and every element \(a\) has a corresponding size \(s(a) \in \mathbb{Z}^+\), \(\sum_{a \in S} s(a) = 2A\). We construct a S-P graph (see Figure 1) as follows.

- \(V = \{v_1, v_2, \ldots, v_{n+1}, u_1, u_2, \ldots, u_n\}\),
- \(E = \{(v_i, v_{i+1}) | i = 1, 2, \ldots, n\} \cup \{(v_i, u_i), (v_{i+1}, u_i) | i = 1, 2, \ldots, n\},\)
- For \(e = (v_i, v_{i+1})\), \(c(e) = s(a_i), w(e) = 0.\)
- For \(e = (v_i, u_i)\) or \((v_{i+1}, u_i)\), \(c(e) = 0, w(e) = \frac{1}{2}s(a_i).\)

For this S-P graph, we want to find a shortest path from \(v_1\) to \(v_{n+1}\) such that the total weight of this path is no larger than \(A\).

If we can find a partition of \(S = S_1 \cup S_2\) such that \(\sum_{a \in S_1} s(a) = \sum_{a \in S_2} s(a) = A\), then we have a path from \(v_1\) to \(v_{n+1}\) consist of the edges \((v_i, v_{i+1})\) for \(a_i \in S_1\), and \((v_i, u_i), (u_i, v_{i+1})\) for \(a_i \in S_2\). The total cost and total weight of this path are all equal to \(A\). This path is obviously the shortest path under the restriction of total weight no larger than \(A\).

Conversely, assume that we have find a shortest path from \(v_1\) to \(v_{n+1}\) with the total weight no larger than \(A\). From the construction of the S-P graph, we can get a partition of \(S\) by setting \(S_1 = \{a_i | (v_i, v_{i+1})\) is in the shortest path\}. From above discussion, we know that the weight constrained shortest path problem in S-P graph is NP-hard.

\[\square\]
Decidable Approximations on Generalized and Parameterized Discrete Timed Automata

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Abstract. We consider generalized discrete timed automata with general linear relations over clocks and parameterized constants as clock constraints and with parameterized durations. We look at three approximation techniques (i.e., the $r$-reset-bounded approximation, the $B$-bounded approximation, and the $\langle B, r \rangle$-crossing-bounded approximation), and derive automata-theoretic characterizations of the binary reachability under these approximations. The characterizations allow us to show that the safety analysis problem is decidable for generalized discrete timed automata with unit durations and for deterministic generalized discrete timed automata with parameterized durations. An example specification written in ASTRAL is used to run a number of experiments using one of the approximation techniques.

1 Introduction

As a standard model for analyzing real-time systems, timed automata \cite{3} have received enormous attention during the past decade. A fundamental result in the theory of timed automata shows that region reachability for timed automata is decidable \cite{3}. This result has been very useful in defining various real-time logics, appropriate model checking algorithms and tools \cite{24,16,17,25,27,28,29} for verifying real-time systems (see \cite{1} for a survey).

However, not every real-time system can be modeled as a timed automaton since a complex (not necessarily large) real-time system might contain non-region (e.g., Presburger) clock constraints. However, the “Turing computing” power of such augmented automata prevents automatic verification of properties such as reachability \cite{3,5}. Therefore, decidable approximation techniques are of great interest, since these techniques would provide a user some degree of confidence in analyzing and debugging complex real-time specifications. In contrast to the most direct approximation techniques \cite{5,10,11,12} that bound the number of transitions to a fixed number, the approximation techniques presented in this paper restrict the clock behaviors but do not necessarily bound the number of transition iterations to be finite.
In this paper, we focus on timed automata with integer-valued clocks, i.e., discrete timed automata, and extend them to generalized discrete timed automata by allowing general linear relations over clocks and parameterized constants as clock constraints. Furthermore, the duration of a transition can be a parameterized constant. These generalizations have practical motivation. For example, many complex real-world specifications [6,10,22] written in the real-time specification language ASTRAL [6] use generalized clock constraints and parameterized durations in almost every specification. Therefore, the results presented in this paper may be useful in implementing an automatic specification debugger for complex real-time specification languages like ASTRAL.

We investigate three approximation techniques in this paper. The $r$-reset-bounded approximation bounds the number of clock resets by a given positive integer $r$ for each clock. The $B$-bounded approximation requires that each clock reset happens before its value exceeds a given (integral) time $B$. Combining these two, the $⟨B,r⟩$-crossing-bounded approximation requires that for each clock, there are at most $r$ times that the clock resets after its value exceeds $B$. Given an approximation technique, we will focus on the binary reachability characterization of a generalized discrete timed automaton. Binary reachability has recently been proposed and investigated for a number of timed systems [7,9,14,15], which makes it possible to reason about “non-region” properties for these systems. We first show that, when a generalized discrete timed automaton has unit duration, the binary reachability under any one of the three approximations is Presburger. Then by considering a generalized discrete timed automaton with parameterized durations, we show that the binary reachability under any one of the three approximations has a 2DCM-padding when the machine is deterministic. Specifically, we show that the “padded language” for binary reachability can be accepted by a deterministic two-way counter machine with one reversal-bounded counter [20]. The case for nondeterministic generalized discrete timed automata is open. These are good characterizations in the sense that the validity of Presburger formulas as well as the emptiness problem for these counter machines are decidable. This allows us to establish, in principle, decidable results for the (Presburger) safety analysis problem for generalized discrete timed automata under the approximations. The 2DCM-padding characterization is particularly interesting, since binary reachability is not necessarily semilinear.

Due to space limitation, all the proofs are omitted. For a complete exposition see [13].

# Preliminaries

Let $V$ be a finite set of variables over the integers. An atomic linear relation on $V$ is defined as $\sum_{v \in V} a_v v < b$, where $a_v$ and $b$ are integers. A linear relation on $V$ is constructed from a finite number of atomic linear relations using ¬ and ∧. $\mathcal{L}_V$ denotes the set of all linear relations on $V$. An atomic Presburger relation on $V$ is either an atomic linear relation on $V$ or a linear congruence $\sum_{v \in V} a_v v = b \mod d$, where $a_v$, $b$ and $d$ are integers. A Presburger formula can always be constructed
from atomic Presburger relations using $\neg$ and $\land$. Let $N$ be the set of integers with $N^+$ denoting the set of nonnegative integers. A clock is a variable over $N^+$. A generalized discrete timed automaton $A$ is a tuple $(S, C, X, E)$ where $S$ is a finite set of (control) states, $X$ is a finite set of clocks, $C$ is a finite set of parameterized constants, and $E \subseteq S \times (C \cup \{0,1\}) \times 2^X \times L_{X \cup C} \times S$ is a finite set of edges. Each edge $(s, d, \lambda, l, s')$ denotes a transition (or an edge) from $s$ to $s'$ with enabling condition $l \in L_{X \cup C}$. $d \in C \cup \{0,1\}$ is a parameterized constant indicating the duration of this transition. $\lambda \subseteq X$ is the set clocks that are reset as a result of this transition. When $\lambda$ is empty, the duration, which is a parameterized constant in $C$ or integer constant 1, must be positive. Clock resets take no time. Thus, when $\lambda$ is not empty, the duration $d$ on this edge is simply 0.

The semantics is defined as follows. $\alpha \in S \times (N^+) \times (N^+)^X$ is called a configuration with $\alpha_q$ being the state under this configuration. $\alpha_{x_i}$ and $\alpha_{c_j}$ denote the value of the clock $x_i$ and the value of the parameterized constant $c_j$, respectively. Note that each clock and parameterized constant are nonnegative, i.e., in $N^+$. $\alpha \rightarrow (s, d, \lambda, l, s')$ $\alpha'$ denotes a one-step transition along an edge in $A$ satisfying

- Constant values do not change, i.e., for each $c \in C$, $\alpha_c = \alpha'_c$.
- The state $s$ is set to a new location $s'$, i.e., $\alpha_q = s, \alpha'_q = s'$.
- Each clock changes according to the edge given. When there is no clock reset on this edge, i.e., $\lambda = \emptyset$, all the clocks synchronously progress by the amount of the duration, i.e., for each $x \in X$, $\alpha'_x = \alpha_x + \alpha_d$. In this case, the duration is positive, i.e., $\alpha_d > 0$. When $\lambda \neq \emptyset$, clocks in $\lambda$ reset to 0 and all the other clocks are unchanged. Thus, clock resets take no time. That is, for each $x \in \lambda$, $\alpha'_x = 0$ and for each $x \notin \lambda$, $\alpha'_x = \alpha_x$.
- The enabling condition is satisfied. That is, $l(\alpha)$ is true.

We simply write $\alpha \rightarrow \alpha'$ if $\alpha$ can reach $\alpha'$ by a one-step transition. $A$ is deterministic if for any configuration $\alpha$ there is at most one $\beta$ such that $\alpha \rightarrow \beta$. A path $\alpha_0 \cdots \alpha_k$ satisfies $\alpha_i \rightarrow \alpha_{i+1}$ for each $i$. Write $\alpha \sim_A \beta$ if $\alpha$ reaches $\beta$ through a path. In the definition of $A$, there is no input. This is because the input is always one-way for timed automata and, therefore, input symbols can be built into the control states. $A$ has parameterized durations. If we restrict each duration on an edge without clock resets to be 1, then $A$ is called a generalized discrete timed automaton with unit durations.

When $A$ with unit durations contains no parameterized constants and enabling conditions are clock constraints or clock regions in the form of Boolean combinations of $x - y \# c$ and $x \# c$, where $c$ is an integer, $x$ and $y$ are clocks and $\# \in \{<, =, >\}$, $A$ is equivalent to the standard timed automata (with integer-valued clocks) [3]. On the other hand, when $A$ with unit durations contains enabling conditions only in the form of Boolean combinations of $x \# c$, where $c$ is a parameterized constant or an integer, $x$ is a clock (when $c$ is a parameterized constant, $x$ is called a parameterized clock), $A$ is a parameterized timed automaton (with integer-valued clocks) [5].
There are two kinds of reachability for $\mathcal{A}$. They are state reachability and binary reachability. Assume a state $s_0$ is designated as the initial state of $\mathcal{A}$. A state $s \in S$ is state reachable in $\mathcal{A}$ if there is an initial configuration $\alpha_0$ (whose state is $s_0$ and whose clock values are all 0) and a configuration $\alpha$ with $\alpha_q = s$ such that $\alpha$ is reachable from the initial configuration, i.e., $\alpha_0 \sim_{\mathcal{A}} \alpha$. The state reachability set is the set of all states $s$ such that $s$ is state reachable in $\mathcal{A}$. The state reachability problem of $\mathcal{A}$ is whether a state $s \in S$ is state reachable. The following are some known results about the state reachability problem.

**Theorem 1.** (1). The state reachability problem is decidable for standard timed automata [3]. (2). The state reachability problem is decidable for parameterized timed automata with only one parameterized clock (but can have many unparameterized clocks) [5]. (3). The state reachability problem is undecidable for generalized discrete timed automata.

Actually, Theorem 1 (3) follows from the following special cases.

**Theorem 2.** (1). The state reachability problem is undecidable for standard timed automata when we allow “+” operations in clock constraints, e.g., $x + y - z < 5$ [3]. (2). The state reachability problem is undecidable for parameterized timed automata (with more than 2 parameterized clocks) [5].

On the other hand, binary reachability is the set of all configuration pairs $\langle \alpha, \beta \rangle$ such that $\alpha \sim_{\mathcal{A}} \beta$ (we use $\sim_{\mathcal{A}}$ in the paper). Characterizations of the binary reachability of timed automata have recently been established.

**Theorem 3.** (1). The binary reachability of timed automata with real valued clocks is definable in the additive theory over reals and integers [7,9]. (2). The binary reachability of timed automata with integer valued clocks is definable by a Presburger formula [7,15].

Characterizations of binary reachability help us to reason about non-region properties of timed systems. For instance, consider the following property: “for every configuration $\alpha$ there exists a configuration $\beta$ such that $\alpha \sim_{\mathcal{A}} \beta$ and the clock $x_1$ in $\beta$ is the sum of clocks $x_1$ and $x_2$ in $\alpha$.” Though constraint $\beta_{x_1} = \alpha_{x_1} + \alpha_{x_2}$ is not in the form of a clock region, this property can be automatically verified for timed automata [7,9,15].

However, for generalized discrete timed automata, the binary reachability $\sim_{\mathcal{A}}$ is too strong to have an interesting characterization. In particular, even the membership problem for binary reachability, i.e., deciding whether two given configurations $\alpha$ and $\beta$ satisfy $\alpha \sim_{\mathcal{A}} \beta$ for a generalized discrete timed automaton $\mathcal{A}$, is undecidable. This follows from the fact that a two-counter machine can be simulated by a generalized discrete timed automaton $\mathcal{A}$, as shown in the the proofs of Theorem 2 (1) (2) [3,5]. Thus, the membership problem can be reduced to the halting problem for two-counter machines, which is undecidable.

**Theorem 4.** The membership problem for binary reachability is undecidable for generalized discrete timed automata.
This undecidability result for generalized discrete timed automata leads us to consider the following three approximations of $\sim_A$. Let $r$ and $B$ be fixed positive integers. The first approximation is $r$-reset-bounded reachability. A path $\alpha^0\alpha^1\cdots\alpha^k$ is called $r$-reset-bounded if each clock resets at most $r$ times. Write $\alpha \sim^r_A \beta$ if $\alpha$ reaches $\beta$ through an $r$-reset-bounded path. The second approximation is $B$-bounded reachability. A path $\alpha^0\alpha^1\cdots\alpha^k$ is called $B$-bounded if for each $j < k$, each $x_i \in X$, $|\alpha^j_i - \alpha^{j+1}_i| < B$. Write $\alpha \sim^B_A \beta$ if $\alpha$ reaches $\beta$ through a $B$-bounded path. The third approximation is $\langle B, r \rangle$-crossing-bounded reachability. A path $\alpha^0\alpha^1\cdots\alpha^k$ is called $\langle B, r \rangle$-crossing-bounded if there are at most $r$ many $i$’s such that $|\alpha^j_i - \alpha^{j+1}_i| \geq B$. Write $\alpha \sim^{\langle B, r \rangle}_A \beta$ if $\alpha$ reaches $\beta$ through a $\langle B, r \rangle$-crossing-bounded path.

The main results in this paper show that the three approximations of binary reachability $\sim_A$ have decidable characterizations, i.e., they can be accepted by a class of machines with a decidable emptiness problem. Before we proceed to show the results, some further definitions are needed.

A nondeterministic multicounter machine (NCM) is a nondeterministic machine with a finite number of states, a one-way input tape, and a finite number of integer counters. Each counter can be incremented by 1, decremented by 1, or stay unchanged. These counter assignments are called standard assignments. In addition, a counter can be tested against an integer constant. These tests are called standard tests. An NCM is reversal-bounded if each counter is reversal-bounded (i.e., it changes mode between nondecreasing and nonincreasing for some bounded number of times).

A tuple of integers can be encoded as a string by concatenating the unary representations of each integer with a separator. In this way, the binary reachability $\sim_A$ can be treated as a language $\{[\alpha]\#[\beta] : \alpha \sim_A \beta\}$, where $[\alpha]$ and $[\beta]$ are string encodings of configurations $\alpha$ and $\beta$ separated by a delimiter “#”. Obviously, same encoding applies to $\sim_M$, the binary reachability of an NCM $M$.

When an NCM is reversal-bounded, the emptiness problem is decidable. In fact, we have a stronger characterization.

**Theorem 5.** A set of $n$-tuples of integers is definable by a Presburger formula iff it can be accepted by a reversal-bounded deterministic multicounter machine [19].

The machines defined above, when used as language recognizers, have a one-way input tape. Suppose a two-way input is used instead. Let $2\text{DCM}(c, r)$ denote the class of deterministic machines with a two-way input tape and $c$ $r$-reversal-bounded counters. Then the emptiness problem for $2\text{DCM}(c, r)$ when $c \geq 2$ and $r \geq 1$ is undecidable [19]. An interesting special case is when $c = 1$, i.e., there is only one counter. A language is $2\text{DCM}$-recognizable if it can be accepted by a $2\text{DCM}(1, r)$.

**Theorem 6.** The emptiness problem for $2\text{DCM}$-recognizable languages is decidable [20].
It is still open whether Theorem 6 holds for nondeterministic machines. That is, whether the emptiness problem for 2NCM(1,r), which is a nondeterministic r-reversal-bounded one counter machine with a two-way input tape, is decidable.

Given a generalized discrete timed automaton $A$, consider a subset of configuration pairs, $R_A \subseteq S \times (N^+)^{|C|} \times (N^+)^{|X|} \times S \times (N^+)^{|C|} \times (N^+)^{|X|}$. One can look at $R_A$ as some sort of reachability relation. For a given $A$, if a 2DCM(1, r) $M_A$ can be effectively constructed such that for every $w$, $w$ is in $R_A$ iff there exists a $w'$ such that $M_A$ accepts $w#w'$, then we say that $R_A$ has a 2DCM-padding. From Theorem 6 it is routine to show the following lemma.

Lemma 1. (1). The emptiness problem for $\{R_A\}_A$ having 2DCM-paddings is decidable. (2). If $R_A^1$ and $R_A^2$ have 2DCM-paddings, then so do the join $R_A^1 \cup R_A^2$ and the composition $R_A^1 \circ R_A^2$.

3 Main Results

Denote $\sim_{0}^{A}$ to be the binary reachability of a generalized discrete timed automaton $A$ through a path without clock resets. Note that $A$ itself can be considered as a nondeterministic multicounter machine. However, tests in $A$, which are linear relations on clocks and parameterized constants, are not standard. Assignments in $A$ in the form of $x := x + d$ with $d$ a parameterized constant are also not standard. But, if $A$ has only unit durations, the assignments are standard except when a clock reset occurs, i.e., $x := 0$. The following result follows our recent results on strong-reversal-bounded NCMs [21].

Lemma 2. Suppose $A$ is a generalized discrete timed automaton with unit durations. Then $\sim_{0}^{A}$ is Presburger.

The following theorem gives a characterization of the three approximations of $\sim_{A}$ when $A$ has unit durations. The proof cuts a reachability path of $A$ into a finite number of phases and each phase can be further characterized by $\sim_{0}^{A}$ using Lemma 2.

Theorem 7. Suppose $A$ is a generalized discrete timed automaton with unit durations. Then $\sim_{r}^{A}$, $\sim_{B}^{A}$ and $\sim_{(B,r)}^{A}$ are Presburger.

The case for $A$ with parameterized durations is more complicated. In principle, $A$ with parameterized durations can be simulated by an $A'$ with unit durations. This is done by simply introducing a new clock $z$ to test whether the parameterized duration of a transition is reached, and after the transition is fired $z$ is reset to 0. However, the three approximations on $A$ are not equivalent to those on $A'$. The reason is as follows. Consider $\sim_{0}^{A}$, the 0-reset-bounded approximation of $A$. $\alpha \sim_{0}^{A} \beta$ if $\alpha$ can reach $\beta$ through a path without clock resets. But for $A'$ each transition with a parameterized duration causes clock $z$ to reset. Thus, a path in $A'$ witnessing $\alpha \sim_{0}^{A} \beta$ could have an unbounded number of clock resets. Therefore, for $A$ with parameterized durations, Theorem 7 is not applicable.
In the following, we consider a generalized discrete timed automaton $A$ (with parameterized durations). Currently, we cannot show a decidable characterization of $\sim^0_A$, since we find it is related to the emptiness problem of $2\text{NCM}(1,r)$, which is still open. However, by restricting $A$ to be deterministic, the following results can be established. The proof uses the fact that the disjunction of the atomic linear relations that appear in all the enabling conditions of $A$ is equivalent to a union of convex linear relations.

**Lemma 3.** $\sim^0_A$ has a $2\text{DCM}$-padding for a deterministic generalized discrete timed automaton $A$.

Again, by using Lemma 3 and the idea of cutting a reachability path of $A$ into phases, the following theorem gives a characterization of the three approximations of $\sim_A$.

**Theorem 8.** Suppose $A$ is a deterministic discrete timed automaton. Then, $\sim^r_A$, $\sim^B_A$ and $\sim^{(B,r)}_A$ have $2\text{DCM}$-paddings.

### 4 Verification of Safety Properties

Consider $P$ and $I$, two sets of configurations of a generalized discrete timed automaton $A$ definable by Presburger formulas. If, starting from a configuration in $I$, $A$ can only reach configurations in $P$, then $P$ is a safety property with respect to the initial condition $I$. The following is an example:

“starting from a configuration satisfying $x_1 - x_2 + x_3 - x_4 > c + d$, $A$ cannot reach a configuration satisfying $2x_1 + x_2 < c \land x_3 - 3x_4 > 2d - c$.”

The safety analysis problem is to determine whether $P$ is a safety property with respect to the initial condition $I$. The following two theorems follow from Theorem 7, Theorem 8, and Lemma 1.

**Theorem 9.** The safety analysis problem is decidable for generalized discrete timed automata with unit durations and with any one of the following approximations: $r$-reset-boundedness, $B$-boundedness and $(B,r)$-crossing-boundedness.

**Theorem 10.** The safety analysis problem is decidable for deterministic generalized discrete timed automata with any one of the following approximations: $r$-reset-boundedness, $B$-boundedness and $(B,r)$-crossing-boundedness.

**Remark:** Theorem 10 can be strengthened. The class of languages accepted by deterministic two-way counter machines with one reversal-bounded counter is closed under Boolean operations [20]. It follows that Theorem 10 remains valid even if the sets of configurations $P$ (property) and $I$ (initial condition) are sets accepted by these machines.

It is desirable to consider the decidability of the safety analysis problem for generalized discrete timed automata under some special form but without using the approximations. One such form is parameterized timed automata with
only one parameterized clock. As stated in Theorem 1(2), the state reachability problem is decidable. However, surprisingly, the safety analysis problem, i.e.,

“Deciding whether $P$ is a safety property with respect to the initial condition $I$ for a parameterized timed automaton with only one parameterized clock, where both $P$ and $I$ are definable by Presburger formulas”
is still open. This problem is closely related to the open problem of the decidability of the emptiness problem for 2NCM($1,r$). It is also worthwhile to point out that the characterizations of the binary reachability of deterministic generalized discrete timed automata under the approximations, as shown in Theorem 8, are not necessarily Presburger. In fact, 2DCM($1,r$) can accept a class of nonlinear languages [20].

5 A Verification Example

In practice, allowing parameterized clock constraints and durations makes it possible to specify more complex real-time systems. In this section, we take an example specification [23] of the railroad crossing benchmark [18], which is written in ASTRAL [6]. The specification specifies a system consisting of a set of railroad tracks that intersect a street where cars may cross the tracks. A gate is located at the crossing to prevent cars from crossing the tracks when a train is near. A sensor on each track detects the arrival of trains on that track. The critical requirement of the system is that whenever a train is in the crossing the gate must be down, and when no train has been in between the sensors and the crossing for a reasonable amount of time, the gate must be up. The complete ASTRAL specification of the railroad crossing system, which was written by Paul Kolano, can be found at http://www.cs.ucsb.edu/~dang.

The ASTRAL specification includes a global specification and two process specifications: one is Gate and the other is Sensor. A transition system is specified inside each process specification along with local assumptions and local properties. ASTRAL adopts a modularized view of the specified system at the level of correctness proofs: the global properties in the global specification can be verified by using global assumptions and local properties (instead of the actual transition behaviors) of each process instance; local properties of a process instance can be verified by using the local assumptions, the local imported variable clause and the transition system of the instance, without looking at the transition behaviors of the other process instances (A reader need not worry about the possibility of circular proofs. The ASTRAL proof theory is sound, see [8] for details).

We take the Sensor process specification to see how parameterized durations and parameterized clock constraints are used in ASTRAL. Sensor reports whether a train is beyond the crossing, which is indicated by a Boolean variable train_in_R. The process specification has only two transitions enter_R and exit_R, which have parameterized durations enter_dur and exit_dur, respectively. Transition enter_R changes train_in_R from FALSE to TRUE dur, respectively. Transition enter_R changes train_in_R from FALSE to TRUE with duration enter_dur. Transition exit_R sets train_in_R back to FALSE after the slowest train moves out of the crossing. That is,
TRANSITION exit_I
ENTRY [TIME: exit_dur]
\text{train\_in\_R} \\
& \text{now - Start ( enter\_R ) } \geq \text{RIImin - exit\_dur}
EXIT
\text{train\_in\_R} = \text{FALSE},

where \text{now} indicates the current time, \text{Start(enter\_R)} is the most recent start time of transition \text{enter\_R}, and \text{RIImin} is a parameterized constant indicating the time for the slowest train to move out of the crossing. One of the two transitions is nondeterministically chosen to fire as long as the ENTRY condition of the chosen transition is satisfied. However, if neither of them is fireable, the process idles: \text{now} progresses by one time unit and \text{train\_in\_R} does not change.

But, according to this semantics, transition \text{enter\_R} must fire immediately whenever transition \text{exit\_R} completes. This is not the intended behavior of \text{Sensor}. In fact, it is desirable that \text{enter\_R} fires only when a train actually comes. But \text{enter\_R} specified as above only tells what happened (set \text{train\_in\_R} to \text{TRUE}) when a train comes, instead of the time when a train comes. The pattern of a train’s arrival is controlled by the environment of the process. In this specification, transition \text{enter\_R} is declared as exported. That is, it must be called by the environment in order to fire. In \text{Sensor}, the environment is specified by the environment clause, which states that two consecutive calls of \text{enter\_R} must be separated by at least \text{RIImin} many time units. The safety property of the process is specified as a \text{schedule}, which has two conjuncts. The first conjunct of the schedule says that a train will be sensed within \text{enter\_dur} many time units after a call of transition \text{enter\_R} is placed. The second conjunct of the schedule says that the sensor will be reset when the slowest train is beyond the crossing. It is assumed that initially \text{now} is 0 and \text{train\_in\_R} is \text{FALSE}.

We manually translated \text{Sensor} into a generalized discrete timed automaton and computed the transitive closure of the one-step transition of the automaton using the Omega Library [26], which is a tool to manipulate Presburger formulas. Experiments were run on a Sun workstation with 4 CPUs and 256M real memory and 512M swap memory. Unfortunately, the closure, even when the durations (\text{enter\_dur} and \text{exit\_dur}) were set to 1, could not be computed. Then, we used the \text{B}-bounded approximation on the automaton with \text{B} = 3. This time, the binary reachability \sim{B} can be computed in about one minute of CPU time and using 170M memory. But the other two approximation approaches were still too expensive to calculate.

6 Conclusions and Future Work

We studied generalized discrete timed automata with general linear relations over clocks and parameterized constants as clock constraints and with parameterized durations. We focused on three approximation techniques and automata-theoretic characterizations of binary reachability under these approximations. The characterizations allow us to show that the safety analysis problem is decidable with respect to generalized discrete timed automata with unit durations,
and deterministic generalized discrete timed automata with parameterized durations (modulo the approximations). We used an example specification written in ASTRAL to run a number of experiments using one of the approximation techniques. The results of the experiments show that further improvements to the approximations have to be developed, since currently they are not practical for large specifications.

For future work, we want to investigate how the approximation techniques proposed in this paper can be combined with existing image-approximation techniques [12] in debugging infinite state systems. Solutions to this problem would lead to an implementation of an effective specification debugger for large real-time specifications. Another research issue is how to extend the results in this paper to the case of generalized timed automata with dense clocks. Recent ideas used in [9] may provide some insights.

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Multiplicative Adaptive Algorithms for User Preference Retrieval

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Abstract. In contrast to the adoption of linear additive query updating techniques in existing popular algorithms for user preference retrieval, in this paper we design two types of algorithms, the multiplicative adaptive query expansion algorithm MA and the multiplicative adaptive gradient search algorithm MG, both of which use multiplicative query expansion strategies to adaptively improve the query vector. We prove that algorithm MA has a substantially better mistake bound than the Rocchio’s and the Perceptron algorithms in learning a user preference relation determined by a linear classifier with a small number of non-zero coefficients over the real-valued vector space $[0,1]^n$. We also show that algorithm MG boosts the usefulness of an index term exponentially, while the gradient descent procedure does so linearly. Our work also generalize the algorithm Winnow in the following aspects: various updating functions may be used; multiplicative updating for a weight is dependent on the value of the corresponding index term, which is more realistic and applicable to real-valued vector space; and finally, a number of documents which may or may not be counterexamples to the algorithm’s current classification are allowed. Practical implementations of algorithms MA and MG have been underway in the next stage development of our intelligent web search tools.

1 Vector Space and User Preference

Let $\mathcal{R}$ be the set of all real values, and let $\mathcal{R}^+$ be the set of all non-negative real values. Let $n$ be a positive integer. In the vector space model in information retrieval [11], a collection of $n$ indexing terms $T_1, T_2, \ldots, T_n$ are used to represent documents and queries. Each document $d$ is represented as a vector $d = (d_1, \ldots, d_n)$ such that for any $i$, $1 \leq i \leq n$, the $i$-th component of $d$ is used to determine the relevance (or weight) of the $i$-th term $T_i$ in the document. Because a document vector can be normalized, without loss of generality we only consider the real-valued vector space $[0,1]^n$ in this paper. Given any two vectors $x = (x_1, \ldots, x_n), y = (y_1, \ldots, y_n) \in [0,1]^n$ (or $\mathcal{R}^n$), we use $x \cdot y$ to denote their inner product $x_1y_1 + \cdots + x_ny_n$.

Let $\mathcal{D}$ be a collection of documents. Given any two documents in $\mathcal{D}$, we assume that a user would prefer one to other or regard both as being equivalent.
with respect to her information needs (or search queries). In other words, user preference of documents in \( D \) defines a preference relation \( \prec \) over \( D \) as follows:

\[
\forall d, d' \in D, \quad d \prec d' \iff \text{the user prefers } d' \text{ to } d.
\]

It has been shown in [1] that if a user preference relation \( \prec \) is a weak order satisfying some additional conditions then it can be represented by a linear classifier. That is, there is a query vector \( q = (q_1, \ldots, q_n) \in \mathbb{R}^n \) such that

\[
\forall d, d' \in D, \quad d \prec d' \iff q \cdot d < q \cdot d'.
\]

(1)

In general a linear classifier over the vector space \([0, 1]^n\) is a pair of \((q, \theta)\) which classifies any document \( d \) as relevant if \( q \cdot d > \theta \), or irrelevant otherwise, where the query vector \( q \in \mathbb{R}^n \) and the classification threshold \( \theta \in \mathbb{R}^+ \). Recall that \( q \cdot d \) is usually used as the relevance rank or (score) of the document \( d \) with respect to user preference.

A natural way to understand a user preference relation \( \prec \) is document ranking: A user prefers a document \( d \) to a document \( d' \), if and only if she ranks \( d \) higher than \( d' \). When a user has no preference of \( d \) to \( d' \) nor \( d' \) to \( d \), then she is really not interested in how those two documents are actually ranked. Based on such understanding, the following linear acceptable ranking strategy was proposed in [13]:

\[
\forall d, d' \in D, \quad d \prec d' \iff q \cdot d < q \cdot d'.
\]

(2)

where \( q \in \mathbb{R}^n \) is the query vector determined by the user.

Let \( D_r \) be the set of all relevant documents in \( D \) with respect to a user’s information needs (or search queries), and \( D_ir \) the set of all irrelevant documents. If we assume that a user preference relation has a simple structure with only two levels, i.e., one level consisting of all relevant documents and the other consisting of all irrelevant documents. Within the same level, no preference is made between any two documents. Then, finding a user preference relation satisfying the expression (1) is equivalent to the problem of finding a linear classifier \((q, \theta)\) over \([0, 1]^n\) with the property

\[
\forall d \in D, \quad d \in R \iff q \cdot d > \theta,
\]

(3)

where \( q \in \mathbb{R}^n \) is the query (or weight) vector. Similarly, finding a linear acceptable ranking strategy satisfying expression (2) is equivalent to the problem of finding a query vector \( q \in \mathbb{R}^n \) with the property

\[
\forall d, d' \in D, \quad d \in D_{ir} \text{ and } d' \in D_{r} \implies q \cdot d < q \cdot d'.
\]

(4)

The goal of relevance feedback in information retrieval is to identify a user preference relation \( \prec \) with respect to her information needs from documents judged by that user. Since user preference relations vary from different users and may have various unknown representations, it is not easy for an information system to find such relations. The existing popular relevance relevance algorithms
basically use linear additive query expansion methods to find a user preference relation as follows:

- Start with an initial query vector $q_0$.  
- At any step $k \geq 0$, improve the $k$-th query vector $q_k$ to  

$$q_{k+1} = q_k + \alpha_1 d_1 + \cdots + \alpha_s d_s,$$

(5)

where $d_1, \ldots, d_s$ are the documents judged by user, and the updating factors $\alpha_i \in \mathbb{R}$ for $i = 1, \ldots, s$.

One particular and well-known example of relevance feedback is Rocchio’s similarity-based relevance feedback [7]. Depending on how updating factors are used in improving the $k$-th query vector as in expression (5), a variety of relevance feedback algorithms have been designed [11]. A similarity-based relevance feedback algorithm is essentially an adaptive supervised learning algorithm from examples [12]. The goal of the algorithm is to learn some unknown classifier (such as the linear classifier in expression (3)) that is determined by a user’s information needs to classify documents as relevant or irrelevant. The learning is performed by modifying or updating the query vector that serves as the hypothetical representation of the collection of all relevant documents. The technique for updating the query vector is linear additions of the vectors of documents judged by the user. This type of linear additive query updating technique is similar to what used by the Perceptron algorithm [10]. The linear additive query updating technique has some disadvantage: It’s **converging rate** to the unknown target classifier is slow, because it has been proved that the Perceptron algorithm [8] and the Rocchio’s relevance feedback algorithm [5] with any of the four typical similar measures [11] have an $\Omega(n)$ lower bound on their performance of learning in the $n$ dimensional binary vector space. In the real world of web search, a huge number of terms (usually, keywords) are used to index web documents. To make the things even worse, no users will have the patience to try, say, more than 10 iterations of relevance feedback in order to gain some significant search precision increase. This implies that the traditional linear additive query updating method may be too slow to be applicable to web search, and this motivates us to design new and faster query updating methods for user preference retrieval in section 4.

For a user preference with respect to her information needs, for any index term $T$, define  

- $|D| = \text{the total number of documents in the collection } D$  
- $|D_r| = \text{the total number of relevant documents}$  
- $\eta = \text{the number of documents indexed by } T$  
- $\gamma = \text{the number of relevant documents indexed by } T$

A gradient descent procedure has been designed in [13] to find an acceptable ranking strategy satisfying expression (2). The idea of the procedure is to minimize ranking errors through linear additions of $d' - d$ for all pairs of documents $d'$ and $d$ that, according to expression (3), are ranked incorrectly. This strategy is similar to what used by the Perceptron algorithm [6]. When the gradient descent procedure is applied to find an acceptable ranking satisfying expression
(4), it has been shown \[13\] that after the first iteration from a zero initial query vector the procedure weighs an index term $T$ linearly in $\frac{\gamma |D_r| - \eta |D|}{|D|}$, an approximate measure of the usefulness of index term $T$ for distinguishing relevant and irrelevant documents. It has also shown that under certain good index term probability distribution the above usefulness measure for the index term $T$ reaches its expected maximum when $\eta = 0.5|D|$, a justification for choosing mid-frequency terms in indexing \[11\]. In contrast to the linear additive query updating strategy used in the gradient descent procedure \[13\] for minimizing ranking errors, in section 4 we will design a new algorithm with the multiplicative query updating strategy. We will show that after the first iteration from a zero initial query vector our new algorithm weighs an index term $T$ exponentially in $\frac{\gamma |D_r| - \eta |D|}{|D|}$. This means that exponentially large gaps will be generated for index terms with respect to measures of their usefulness. Hence, a document with a good index term will be ranked exponentially higher than one without the good index term, thus more ideal ranking effect will be generated for users.

2 The Multiplicative Adaptive Query Expansion Algorithm

In this section we design a multiplicative query updating technique to identify a user preference relation satisfying expression (1). We believe that linear additive query updating yields some mild improvement on the hypothetical query vector towards the target user preference. We want a query updating technique that can yield dramatic improvements so that the hypothetical query vector can be moved towards the target in a much faster pace. Our idea is that when an index term is judged by the user, its corresponding value in the hypothetical query vector should be boosted by a multiplicative factor that is dependent on the value of the term itself.

**Algorithm MA($q_0, f, \theta$):**

(i) Inputs:

- $q_0$, the non-negative initial query vector
- $f(x) : [0, 1] \rightarrow R^+$, the updating function
- $\theta \geq 0$, the classification threshold

(ii) Set $k = 0$.

(iii) Classify and rank documents with the linear classifier ($q_k, \theta$).

(iv) While (the user judged the relevance of a document $d$) do {

for $i = 1, \ldots, n$, do {

/* $q_k = (q_{1,k}, \ldots, q_{n,k})$, $d = (d_1, \ldots, d_n)$ */

if ($d_i \neq 0$) {

if ($q_k \neq 0$) set $q_{i,k+1} = q_{i,k}$ else set $q_{i,k+1} = 1$ /*

adjustment */

if ($d$ is relevant) /* promotion */

set $q_{i,k+1} = (1 + f(d_i))q_{i,k+1}$

else /* demotion */
In this paper, we are only interested in non-decreasing updating functions $f(x) : [0, 1] \mapsto \mathbb{R}^+$, because we want multiplicative updating for an index term is proportional to the value of the term.

We now analyze the the performance of algorithm MA when it is used to identify a user preference satisfying expression (3), a linear classifier $(\mathbf{q}, 0)$. Here, we consider a zero threshold. We say that algorithm makes a classification error at step $k$ if the user judged a document as a counterexample to the algorithm’s current hypothesis. We estimate the total number of classification errors algorithm MA will make based on the worst-case analysis. We also let at most one counterexample may be provided to the algorithm at each step. From now on to the end of this section we assume that $\mathbf{q}$ is a non-negative query vector with $m$ non-zero components and $\theta > 0$. Define

$$\beta = \min\{q_1, \ldots, q_n\}.$$

**Definition 1.** We say that documents in the collection $\mathcal{D}$ are indexing with respect to a threshold $\delta$, $0 < \delta \leq 1$, if for any document $\mathbf{d} = (d_1, \ldots, d_n) \in \mathcal{D}$ we have either $d_i = 0$ or $\delta \leq d_i$, $1 \leq i \leq n$.

In other words, when a document is indexed with respect to a threshold $\delta$, any index term with a value below the threshold $\delta$ is considered not significant, and hence set to zero. Recall that in the vector space a document and its vector have the equivalent meaning, so we may not distinguish the two concepts.

**Lemma 1.** Assume that documents are indexed with respect to a threshold $\delta$. Let $u$ denote the total number of promotions algorithm MA needs to find the linear classifier $(\mathbf{q}, 0)$. Let $m$ denote the number of non-zero components in $\mathbf{q}$. Then,

$$u \leq \frac{m \log \frac{\theta}{\beta \delta}}{\log(1 + f(\delta))}.$$  

Proof. Without loss of generality, we may further assume that the $m$ non-zero components of $\mathbf{q}$ are $q_1, \ldots, q_m$. When a promotion occurs at step $k$, a relevant document $\mathbf{d}$ is given to algorithm as a counterexample its classification. Because of document indexing with respect to the threshold $\delta$, there is some $i$ with $1 \leq i \leq m$ such that $d_i \geq \delta$. This means that the $i$-th component $q_{i,k}$ of the query vector $\mathbf{q}_k$ will be promoted to

$$q_{i,k+1} = (1 + f(d_i))q_{i,k} > (1 + f(\delta))q_{i,k},$$  

(v) If the user has not judged any document in the $k$-th step, then stop. Otherwise, let $k = k + 1$ and go to step (iv).

/* The end of the algorithm MA */
because $f$ is non-decreasing. Since $q_{i,k}$ will never be demoted, it follows from expression (6) that $q_{i,k}$ can be promoted at most

$$\frac{\log \frac{\theta}{\beta \delta}}{\log(1 + f(\delta))}$$

(7) times. Since each promotion yields a promotion for at least one $q_{i,k}$ for $1 \leq i \leq m$, the total number of promotions $u$ is at most $m$ times the value given in expression (7).

**Theorem 1.** Assume that documents are indexed with respect to a threshold $\delta$. Let $T$ denote the total number of classification errors algorithm MA makes in order to find the linear classifier $(q, 0)$ over the real-valued vector space $[0, 1]^n$. Let $m$ denote the number of non-zero components in $q$. Then,

$$T \leq \left[(1 + f(1))(n - m) + \gamma \right](1 + f(\delta))\frac{\delta}{f(\delta)} + \left(\frac{f(1)(1 + f(\delta))}{f(\delta)} + 1\right) \frac{m \log \frac{\theta}{\beta \delta}}{\log(1 + f(\delta))}$$

(Hence, if $\theta = \frac{n}{k}$ is chosen, $T = O(k \log n)$.)

**Proof.** Without loss of generality, we may assume that the $m$ non-zero components of $q$ are $q_1, \ldots, q_m$. We estimate the sum of the weights $\sum_{i=1}^n q_{i,k}$. Let $u$ and $v$ be the number of promotion steps and the number of demotion steps occurred during the learning process, respectively. Let $t_k$ denote the number of zero components in $q_k$ at step $k$. Note that once a component of $q_k$ is promoted to a non-zero value, it will never become zero again. For a promotion at step $k$ with respect to a relevant document $d$ judged by the user, for $i = 1, \ldots, n$, we have

$$q_{i,k+1} = \begin{cases} q_{i,k}, & \text{if } d_i = 0, \\ (1 + f(d_i)), & \text{if } d_i \neq 0 \text{ and } q_{i,k} = 0, \\ (1 + f(d_i))q_{i,k}, & \text{if } d_i \neq 0 \text{ and } q_{i,k} \neq 0. \end{cases}$$

Since a promotion only occurs when

$$q_k \cdot d = \sum_{i=1}^n d_i q_{i,k} = \sum_{d_i \neq 0 \text{ and } q_{i,k} \neq 0} q_{i,k} < \theta,$$

we have

$$\sum_{i=1}^n q_{i,k+1} = \sum_{d_i \neq 0 \text{ and } q_{i,k} = 0} q_{i,k+1} + \sum_{d_i \neq 0 \text{ and } q_{i,k} \neq 0} q_{i,k+1} + \sum_{d_i = 0} q_{i,k+1}$$

$$= \sum_{d_i \neq 0 \text{ and } q_{i,k} = 0} (1 + f(d_i)) + \sum_{d_i \neq 0 \text{ and } q_{i,k} \neq 0} (1 + f(d_i))q_{i,k} + \sum_{x_i = 0} q_{i,k}$$

$$\leq (1 + f(1))t_k + \frac{1 + f(1)}{\delta} \sum_{d_i \neq 0 \text{ and } q_{i,k} \neq 0} \delta q_{i,k} + \sum_{i=1}^n q_{i,k}.$$
\[
\leq (1 + f(1))t_k + \frac{1 + f(1)}{\delta} \sum_{d_i \neq 0 \text{ and } q_{i,k} \neq 0} d_i q_{i,k} + \sum_{i=1}^{n} q_{i,k}.
\]

(8)

For a demotion at step \(k\) with respect to an irrelevant document \(d\) judged by the user, for \(i = 1, \ldots, n\), we have

\[
q_{i,k+1} = q_{i,k} - (1 - \frac{1}{1 + f(d_i)})q_{i,k} \leq q_{i,k} - (1 - \frac{1}{1 + f(\delta)})q_{i,k}.
\]

Since a demotion occurs only when \(\sum_{i=1}^{n} d_i q_{i,k} > \theta\), we have

\[
\sum_{i=1}^{n} q_{i,k+1} \leq \sum_{i=1}^{n} q_{i,k} - (1 - \frac{1}{1 + f(\delta)}) \sum_{i=1}^{n} q_{i,k}
\]

\[
= \sum_{i=1}^{n} q_{i,k} - \frac{f(\delta)}{(1 + f(\delta))\delta} \sum_{i=1}^{n} \delta q_{i,k}
\]

\[
\leq \sum_{i=1}^{n} q_{i,k} - \frac{f(\delta)}{(1 + f(\delta))\delta} \sum_{i=1}^{n} d_i q_{i,k}
\]

\[
\leq \sum_{i=1}^{n} q_{i,k} - \frac{f(\delta)}{(1 + f(\delta))\delta} \theta
\]

(9)

Let the sum of the initial weights be \(\sigma\). Hence, after \(u\) promotions and \(v\) demotions,

\[
\sum_{i=1}^{n} q_{i,k+1} \leq (1 + f(1)) \sum_{i=1}^{u} t_s + \sum_{i=1}^{n} q_{i,0} + \frac{f(1)\theta u}{\delta} - \frac{f(\delta)\theta v}{(1 + f(\delta))\delta}
\]

\[
\leq (1 + f(1))(n - m) + \sigma + \frac{f(1)\theta u}{\delta} - \frac{f(\delta)\theta v}{(1 + f(\delta))\delta}
\]

Note that at any step the weights are never negative. It follows from the above relation that

\[
v \leq \frac{[(1 + f(1))(n - m) + \sigma](1 + f(\delta))\delta}{f(\delta)\theta} + \frac{f(1)(1 + f(\delta))u}{f(\delta)}.
\]

(10)

It follows from Lemma 3.2 and the above analysis that the total number of promotions and demotions, i.e., the total number of classification errors \(T\), is bounded by

\[
T \leq v + u
\]

\[
\leq \frac{[(1 + f(1))(n - m) + \sigma](1 + f(\delta))\delta}{f(\delta)\theta} + \frac{f(1)(1 + f(\delta))u}{f(\delta)} + u
\]

\[
\leq \frac{[(1 + f(1))(n - m) + \sigma](1 + f(\delta))\delta}{f(\delta)\theta} + \left(\frac{f(1)(1 + f(\delta))}{f(\delta)} + 1\right) \frac{m \log \frac{\theta}{\beta_\delta}}{\log(1 + f(\delta))}
\]

This completes our proof.
3 The Multiplicative Gradient Search Algorithm

In this section, we design algorithm MG for finding a query vector \( q \) satisfying the acceptable ranking strategy condition (2). Algorithm MG uses a multiplicative query updating technique to minimize its ranking errors.

**Algorithm MG** \((q_0, f)\):

(i) **Inputs:**
- \( q_0 \), the non-negative initial query vector
- \( f(x) : [0, 1] \rightarrow \mathbb{R}^+ \), the update function

(ii) Set \( k = 0 \).

(iii) Let \( q_k \) be the query vector at step \( k \). Identify the set of mistakes

\[ \Gamma(q_k) = \{ <d, d'> | d \prec d', q_k \cdot d \geq q_k \cdot d' \} . \]

If \( \Gamma(q_k) = \emptyset \), stop.

(iv) For each pair \( <d, d'> \in \Gamma(q_k) \), do 

\[ \begin{align*}
&\text{for } i = 1, \ldots, n, \text{ do } \{ \\
&\quad \text{if } (d'_i \neq 0) \text{ set } q_{i,k+1} = q_{i,k} \text{ else set } q_{i,k+1} = 1 /* adjustment */ \\
&\quad \text{if } (q_{i,k} \neq 0) \text{ set } q_{i,k+1} = (1 + f(d'_i))q_{i,k+1} / * \text{promotion} */ \\
&\quad \text{else set } q_{i,k+1} = \frac{q_{i,k+1}}{1 + f(d_i)} / * \text{demotion} */ \\
&\text{\} }
\end{align*} \]

(v) Let \( k = k + 1 \) and go to step (iv).

/* The end of the algorithm MG */

**Theorem 2.** Assume that algorithm MG is applied to find an acceptable ranking strategy satisfying condition (4). If one chooses the initial query vector \( q_0 = 0 \), then after the first iteration, for any \( 1 \leq i \leq n \), the weight \( q_{i,1} \) for the \( i \)-th indexing term in \( q_1 \) is

\[ q_{i,1} = \prod_{d \in D_{ir}} \prod_{d' \in D_r} \frac{1 + f(d'_i)}{1 + f(d_i)} . \]

In particular, when a linear updating function \( f(x) = \alpha \) is chosen,

\[ q_{i,1} = (1 + \alpha)^{|D_{ir}|} \left| D_r \right| \left( \frac{\gamma}{\alpha} - \frac{\alpha}{\gamma} \right) . \]

**Proof.** Since the acceptable ranking strategy satisfying condition (4), it follows from \( q_0 = 0 \) that at the first iteration one obtains

\[ \Gamma(q_0) = \{ <d, d'> | d \in D_{ir}, d' \in D_r \} . \]

Note that during the first iteration, for each pair \( <d, d'> \in \Gamma(q_0) \), a promotion performed for \( d' \) and a demotion is performed for \( d \). This implies that for any
After the first iteration the value $q_{i,1}$ with respect to the $i$-th index term is

$$q_{i,1} = \prod_{d \in D_{ir}} \prod_{d' \in D_r} \frac{1 + f(d'_i)}{1 + f(d_i)}.$$ 

When $f$ is a constant function mapping with $f(x) = \alpha$, it easily follows from the above expression

$$q_{i,1} = \prod_{d \in D_{ir}} \prod_{d' \in D_r} \frac{1 + f(d'_i)}{1 + f(d_i)} = \frac{\prod_{d \in D_{ir}} \prod_{d' \in D_r} (1 + f(d'_i))}{\prod_{d \in D_{ir}} \prod_{d' \in D_r} (1 + f(d_i))}$$

$$= \frac{\prod_{d \in D_{ir}} (1 + \alpha)^{|D_{ir}|}}{\prod_{d \in D_r} (1 + \alpha)^{|D_r|}} = \frac{(1 + \alpha)^{|D_{ir}| \gamma}}{(1 + \alpha)^{|D_r|(\eta - \gamma)}}$$

$$= (1 + \alpha)^{|D_r|(\eta - |D_{ir}|) - \gamma |D_{ir}| |D_r| \gamma} = (1 + \alpha)^{|D_r|(\eta - |D_{ir}|) \gamma - |D_{ir}| (\eta - \gamma)}$$

This completes our proof.

4 Concluding Remarks

Our motivations of the work in this paper come from the reality of web search: Web search users usually have no patience to try, say, more than 10 iterations of relevance feedback for some intelligent search system in order to gain certain significant search precision increase. Existing popular algorithms for user preference retrieval have their own beauty and advantages. Because of the adoption of linear additive query updating techniques, when used to identify user preference determined by a linear classifier those algorithms, such as Rocchio’s algorithm and its variants, the Perceptron algorithm and gradient descent procedure, have either a slow converging rate or small boosting on the usefulness of an index term.

In contrast to the adoption of linear additive query updating techniques in the those existing algorithms, we design two types of algorithms, the multiplicative adaptive query expansion algorithm MA and the multiplicative adaptive gradient search algorithm MG, both of which use multiplicative query updating techniques to adaptively improve the query vector. We have analyzed the worst-case performance of algorithms MA and MG. We plan to investigate their average-case performance. We feel that this task is very challenging. We also plan to conduct empirical studies to understand the behaviors of the algorithm MA and GA with real world data sets. Finally, we would like to point out that implementations of algorithms MA and MG have been underway in the next stage development of our intelligent web search tools [2,3,4].
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References

Parametric Scheduling for Network Constraints

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Abstract. The problem of parametric scheduling is concerned with checking whether a job-set is parametrically schedulable, subject to a set of imposed constraints. In real-time scheduling, parameterization of the schedule plays an important role in extending the flexibility of the scheduler, particularly in the presence of variable execution times. It has been shown that the existence of parametric schedules can be determined in polynomial time when the constraints are restricted to those that can be represented by a network, unimodular matrix. In this paper, we extend the class of constraints for which parametric schedules can be determined efficiently to include network constraints, such as weighted sum of completion times.

1 Introduction

Uncertainty in problem parameters is often a feature of real-time scheduling [16]. In the literature, there exist two broad approaches to address uncertainty, viz. stochastic and deterministic. In stochastic scheduling, the goal is to provide probabilistic guarantees that the constraints imposed on the job-set will be met, under the assumption that the non-constant problem parameters belong to a fixed well-understood distribution [11,14]. This approach is not applicable in the case of “hard” real-time systems [16,12,10], where the guarantees have to be absolute i.e. there is no room for error. Hard real-time systems are typically composed of mission-critical tasks, wherein the consequences of failure, i.e. a violation of constraints can be catastrophic. Hard real-time systems call for deterministic approaches to the issue of uncertainty. Some of the common approaches include worst-case assumptions [11], static scheduling [19] and parametric scheduling [5]. Worst-case assumptions regarding execution times (say) run the risk of constraint violation at run-time [17] and hence the strategy is not always correct. Static approaches make extremely conservative assumptions about each constraint in order to determine the existence of a feasible schedule. This approach is correct inasmuch as the goal is to provide a set of start-times that cannot cause constraint violation. However, static scheduling is extremely inflexible and even simple constraint sets (see Section 3) may not have static schedules. Parametric scheduling attempts to combine the guarantees provided by static scheduling with the flexibility of stochastic scheduling to enhance the class of constraints for which feasible schedules exist.
In real-time scheduling literature, it has been shown that polynomial time algorithms exist for determining parametric schedules, when the constraints are restricted to be “standard” \[5,2\]. In this paper we are concerned with the following question: Can parametric schedulability be determined efficiently for non-standard constraints, with at most 2 tasks per constraint? We provide an affirmative answer to the above question by designing a polynomial time algorithm for network constraints.

The rest of this paper is organized as follows: Section \[2\] provides a formal description of the problem under consideration. Section \[3\] discusses the motivation underlying our research while related approaches are detailed in Section \[4\]. Our algorithm and the accompanying analysis are presented in Section \[5\].

2 Statement of Problem

2.1 Job Model

Assume an infinite time-axis divided into windows of length \(L\), starting at time \(t = 0\). These windows are called periods or scheduling windows. There is a set of non-preemptive, ordered jobs, \(J = \{J_1, J_2, \ldots, J_n\}\) that execute in each scheduling window.

2.2 Constraint Model

The constraints on the jobs are described by System (1):

\[
A \cdot [s, e] \leq b, \quad e \in E,
\]

where,

- \(A\) is an \(m \times 2.n\) rational matrix, in which every row represents a network constraint (to be defined below),
- \(E\) is the convex set defined by the axis-parallel hyper-rectangle (aph)

\[
\Upsilon = [l_1, u_1] \times [l_2, u_2] \times \ldots \times [l_n, u_n]
\]

- \(s = [s_1, s_2, \ldots, s_n]\) is the start time vector of the jobs, and
- \(e = [e_1, e_2, \ldots, e_n] \in E\) is the execution time vector of the jobs

The characterization of \(E\) as an aph is intended to model the fact that the execution time \(e_i\) of job \(J_i\) is not a fixed constant, but can assume any value in the pre-specified range \([l_i, u_i]\), depending on factors such as loop-length. In real-time systems such as Maruti, a number of runs of the job-set are carried out to statistically estimate the range \([l_i, u_i]\) \[10\].

Observe that System (1) can be rewritten in the form

\[
G \cdot s + H \cdot e \leq b, \quad e \in E
\]
Definition 1. **Standard Constraint:** A constraint is defined to be standard, if it represents a relative separation relationship between at most two tasks, i.e. matrices $G$ and $H$ are flow graph, unimodular, with the added provision, that the entry $H[i,j]$ must equal $G[i,j]$ if it is non-zero. For instance a constraint of the form: $s_i - s_j \leq -8$, which specifies that Job $J_j$ must start 8 units after job $J_i$ starts is a standard constraint. A detailed description of standard constraints is available in [5]. Standard constraints can be represented by edges of a flow graph [18].

Definition 2. **Network Constraint:** A constraint is said to be a network constraint, if it can be expressed in the form

$$a.s_i + b.s_j \leq c.e_i + d.e_j, a, b, c, d \in \mathbb{Q}$$

i.e. all co-efficients are arbitrary rational numbers.

Network constraints also have a graph structure; the “edge” representing the set of constraints between two jobs $J_i$ and $J_j$ form a polyhedron in the 4 variables $s_i, e_i, s_j, e_j$, with $e_i$ and $e_j$ being universally quantified [6].

2.3 Query Model

Suppose that job $J_a$ has to be dispatched. We assume that the dispatcher has access to the start times $\{s_1, s_2, \ldots, s_{a-1}\}$ and execution times $\{e_1, e_2, \ldots, e_{a-1}\}$ of the jobs $\{J_1, J_2, \ldots, J_{a-1}\}$.

Definition 3. A parametric schedule of an ordered set of jobs, in a scheduling window, is a vector $s = [s_1, s_2, \ldots, s_n]$, where $s_1$ is a rational number and each $s_i, i \neq 1$ is a function of the start time and execution time variables of jobs sequenced prior to job $J_i$, i.e. $\{s_1, e_1, s_2, e_2, \ldots, s_{i-1}, e_{i-1}\}$. Further, this vector should satisfy the constraint system (1) for all execution time vectors $e \in \mathbb{Y}$.

The combination of the Job model, Constraint model and the Query model constitutes a scheduling problem specification within the E-T-C scheduling framework [17].

The Parametric Scheduling problem is concerned with the following two issues:

1. Determining whether the given job-set has a parametric schedule, i.e. a schedule as defined in Definition (3);
2. Computing the start-time of a job in each scheduling window, assuming that
   (a) The parametric schedulability query has been decided affirmatively, and
   (b) The start and execution times of all jobs sequenced before it, are provided.
   
   This corresponds to the online dispatching phase.

The discussion above directs us to the following formulation of the parametric schedulability query:

$$\exists s_1 \ \forall e_1 \in [l_1, u_1] \ \exists s_2 \ \forall e_2 \in [l_2, u_2] \ldots \exists s_n \ \forall e_n \in [l_n, u_n] \ \ A.[s, e] \leq b \ ?$$

(5)
Motivation

The motivation for Parametric Scheduling has been provided in great detail in [5] and [17]. The key issue is that a simple constraint system such as:

\[
\begin{align*}
  s_1 + e_1 &\leq s_2 \\
  s_2 &\leq s_1 + e_1 + 1 \\
  e_1 &\in [3, 5]
\end{align*}
\]

does not have a static schedule i.e. no assignment of rational numbers to the start times can guarantee the meeting of both constraints for all values of \( e_1 \) in the range [3, 5]. Note that the schedule represented by \( s = [0, s_1 + e_1]^T \) is a valid, feasible, albeit parametric schedule.

An interesting line of research is obtaining schedules that satisfy certain optimization criteria. The complexity of Parametric Optimization for general constraints is not known. However, we can approximate optimization functions involving at most two tasks through the use of network constraints. Optimization criteria, formulated as performance metrics arise in various situations including Job-shop, Flow-shop and Machine-Shop [11]. Typical performance metrics are Makespan, Sum of Completion times, Weighted sum of completion times (also called aggregate metrics in the Operations Research literature), Lateness and Tardiness. For instance, the need to minimize the weighted sum of completion times of jobs \( J_1 \) and \( J_2 \) can be approximated through \( w_1.(s_1+e_1)+w_2.(s_2+e_2) \leq k \), for suitably chosen \( k \) [16]. The other performance metrics can be similarly approximated.

Related Work

The concept of Parametric Scheduling was introduced in [12]. In [5], polynomial time algorithms were presented for the case, when the constraints imposed on the job-set are “standard”. [18] argued that “standard” constraints could be represented by a flow, graph (i.e. the matrix \( A \) is network, unimodular). [2] extends the “standard constraint” model to include the case, in which constraints can exist between adjacent windows. In [17], it is shown that the problem can be solved in PSPACE, when the constraints are arbitrary. However, no hardness result is known for this problem [7].

A curious feature of the algorithms proposed in the literature for Parametric Schedulability is that none of them exploit the ordering information in the job-set. In both [5] and [2] polynomial time bounds are derived by observing that the number of relative constraints between any two tasks is bounded, if only strict relative constraints are permitted. In this paper, we explicitly take into account the ordering information to develop two new concepts viz. Constraint Domination and Constraint Orientation, which in turn are used to develop polynomial time algorithms for testing parametric schedulability in arbitrary network constraints.
Linear programs with at most 2 variables per constraint (LI(2)s) have received quite a bit of attention in the Operations Research community. [15] was the first to observe the correspondence between LI(2)s and graphs; in [6], the Fourier-Motzkin elimination procedure was used to provide a strongly polynomial algorithm which to date is the fastest known algorithm for this problem.

5 Algorithms and Complexity

Algorithm (5.1) presents our strategy for testing parametric schedulability.

```
Function PARAMETRIC-SCHEDULER (Y, A, b)
1: for (i = n down to 2) do
2:   ELIM-UNIV-VARIABLE(e_i)
3:   if (CHECK-INCONSISTENCY()) then
4:       return (false)
5:   end if
6:   PRUNE-CONSTRAINTS()
7:   ELIM-EXIST-VARIABLE(s_i)
8:   if (CHECK-INCONSISTENCY()) then
9:       return (false)
10:  end if
11: end for
12: ELIM-UNIV-VARIABLE(e_1)
13: if (a ≤ s_1 ≤ b, a, b ≥ 0) then
14:   Valid Parametric Schedule Exists
15:   return
16: else
17:   No Parametric Schedule Exists
18:   return
19: end if
```

Algorithm 5.1: A Quantifier Elimination Algorithm for determining Parametric Schedulability

Algorithm (5.2) describes the procedure for eliminating the universally quantified execution variable $e_i \in [l_i, u_i]$. The Fourier-Motzkin elimination technique discussed in [20] represents one implementation of ELIM-EXIST-VARIABLE. In general, any polyhedral projection method suffices. In our work, we assume that the Fourier-Motzkin procedure is used to eliminate the existentially quantified (start-time) variables. When a variable (start-time or execution time) is eliminated, inconsistencies and/or redundancies could result. CHECK-INCONSISTENCY() identifies inconsistencies and declares the system to be infeasible, while PRUNE-CONSTRAINTS() identifies redundancies and eliminates them.

A detailed exposition of the Fourier-Motzkin elimination procedure is available in [20].
Function \textsc{Elim-Univ-Variable}(\textit{A, b})

1: Substitute $e_i = l_i$ in each constraint that can be written in the form $e_i \geq ()$
2: Substitute $e_i = u_i$ in each constraint that can be written in the form $e_i \leq ()$

Algorithm 5.2: Eliminating Universally Quantified variable $e_i \in [l_i, u_i]$

The correctness of Algorithm (5.2) has been argued in [5], while the correctness of the Fourier-Motzkin procedure is discussed in [13].

We point out that Algorithm (5.1) is similar to the ones outlined in the literature. The difference is the implementation of procedures \textsc{Prune-Constraints}() and the analysis provided in §5.1

5.1 Analysis

Observe that the procedure \textsc{Elim-Univ-Variable}() does not increase the number of constraints. However, \textsc{Elim-Exist-Variable}() has the potential to increase the number of constraints substantially. Assuming that the Fourier-Motzkin elimination algorithm is used, the elimination of $k$ start-time variables, could result in the creation of as many as $m^{2k}$ constraints. One such pathological example is provided in [13]. In [5] and [18], it was pointed out that “standard” constraints are closed under Fourier-Motzkin elimination i.e. the elimination of an existential variable results in the set of constraints staying standard. Using the notation, in [18], this corresponds to saying that contracting a vertex of the flow graph representing the constraint set, does not destroy the its graph structure. Since a graph has at most $O(n^2)$ edges at all times, the polynomiality of the algorithm for standard constraints follows.

In our case though, there is no obvious way to either represent the set of constraints or bound their number under Fourier-Motzkin elimination, since in addition to relative constraints, we also have sum constraints, as discussed in Section §3. We make the following observation:

Observation 51 Network constraints are closed under Fourier-Motzkin elimination, using Algorithm (5.1).

Observation (51) follows from the fact an existential variable i.e. a start time variable is eliminated only after the corresponding execution time variable has been eliminated. Consequently, its elimination results in a network constraint between two other jobs. For instance consider the following constraint set:

1. $s_3 \leq s_1 + e_1 + 14$
2. $s_1 + s_3 \leq 22$
3. $s_2 + 22 \leq s_3 + e_3$
4. $e_3 \in [3, 5]$

The elimination of $e_3$ results in:

1. $s_3 \leq s_1 + e_1 + 14$
2. $s_1 + s_3 \leq 22$
3. $s_2 + 19 \leq s_3$
The elimination of $s_3$ (by pairing off constraints in which $s_3$ occurs with opposite polarity) gives rise to the following set of constraints:

1. $s_2 + 19 \leq s_1 + e_1 + 14$;
2. $s_2 + 19 \leq 22 - s_1$;

The key point is that the network structure is preserved, under the elimination. Observe that if $e_3$ were not eliminated, prior to eliminating $s_3$ the closure claim of Observation \([51]\) no longer holds.

Let $S_{ij}$ denote the set of constraints between the two jobs $J_i$ and $J_j$ ($i < j$). We now present an informal overview on the nature of constraints $l \in S_{ij}$. Informally, a relative constraint between two jobs either specifies increased separation between them or decreased separation. For instance, the constraint $s_1 + 8 \leq s_2$ specifies increased separation, while the constraint $s_2 \leq s_1 + 17$ specifies decreased separation. An aggregate constraint either pushes the jobs to the left or to the right. For instance the constraint $s_1 + s_2 \leq 7$ pushes jobs $J_1$ and $J_2$ towards the left, i.e. towards 0, while the constraint $s_3 + s_4 \geq 8$ pushes jobs $J_3$ and $J_4$ towards the right, i.e. towards $L$.

To proceed with our analysis, we need the following definitions. We associate a type with every constraint, specifying whether it a relative constraint or an aggregate constraint.

**Definition 4.** Sum (Aggregate) constraint: A network constraint of the form $a.s_i + b.s_j \leq ()$ is said to be a sum constraint if both $a$ and $b$ have the same sign.

For instance, $s_1 + s_2 \leq 7$ and $-3.s_1 - 4.s_2 \leq -9$ are sum constraints.

**Definition 5.** Difference (Relative) constraint: A network constraint of the form $a.s_i + b.s_j \leq ()$ is said to be a difference constraint, if $a$ and $b$ have opposite signs.

For instance, the constraint $s_1 - s_2 \leq -4$ is a difference constraint.

**Definition 6.** Constraint orientation (Right): A constraint $l \in S_{ij}$ is said to have a right orientation if it specifies increased separation between $J_i$ and $J_j$ (in case of difference constraints), or pushes both jobs to the right (in case of sum constraints).

**Definition 7.** Constraint orientation (Left): A constraint $l \in S_{ij}$ is said to have a left orientation if it specifies decreased separation between $J_i$ and $J_j$ (in case of difference constraints), or it pushes both jobs to the left (in case of sum constraints).

For instance, the constraint $s_1 + e_1 + 4 \leq s_2$ specifies that Job $J_2$ should start at least 4 units after $J_1$ finishes. Since it specifies increased separation, it is has a right orientation. Likewise, the constraint $s_1 + s_3 \leq 12$ requires that $J_1$ and $J_3$ move leftward and hence has a left orientation. Using the flow graph terminology in \([15]\), a forward edge in the constraint graph has a right orientation and a backward edge has a left orientation.
Every constraint $l \in S_{ij}, \forall i, j = 1, \ldots, n$ has an orientation, on account of the total ordering on the job-set. Thus a network constraint between job $J_1$ and $J_5$ (say) has the effect of either drawing them together or pushing them apart. This is not true, if there is no ordering on the job-set. The total ordering on the start time variables implies that all these variables have an interpretation on the same real axis $[0, L]$. In the absence of the total order, each variable has to be interpreted on its own axis. Also see [6].

**Definition 8.** **Comparable constraints:** Two constraints $l_1, l_2 \in S_{ij}$ are said to be comparable if they have the same orientation and type.

Note that only constraints between the same set of jobs are comparable, i.e. a constraint $l \in S_{13}$ and a constraint $l' \in S_{12}$ are not comparable, regardless of their orientation and type.

Constraint comparability is an *equivalence relation* partitioning the set of constraints between two jobs $S_{ij}$ into the following four categories:

1. Difference Constraint with left orientation ($S^1_{ij}$);
2. Difference Constraint with right orientation ($S^2_{ij}$).
3. Sum Constraint with left orientation ($S^3_{ij}$);
4. Sum Constraint with right orientation ($S^4_{ij}$);

**Definition 9.** **Constraint domination:** A constraint $l_1$ is said to dominate another constraint $l_2$ if and only if they are comparable and $l_1 \Rightarrow l_2$ i.e. $l_2$ is satisfied, whenever $l_1$ is satisfied ($l_1, l_2 \in S_{ij}$).

In some sense, the domination relationship attempts to identify constraints which are redundant. For instance, $s_1 - s_2 \leq -4$ is clearly dominated by $s_1 - s_2 \leq -8$, since if the latter constraint is satisfied, the former is trivially met. The interesting case is the comparison between constraints in which there exist execution time variables. Consider the two comparable constraints:

1. $l_1 : s_1 - s_2 \leq -4$
2. $l_2 : s_1 - s_2 \leq -e_1; e_1 \in [3, 5]$.

Observe that $l_2$ still dominates $l_1$; the parametric schedulability query is:

$$\exists s_1 \forall e_1 \in [3, 5] \exists s_2 \ldots A. \langle s, e \rangle \leq b?$$

Since the query is true for all values of $e_1$ in the range $[3, 5]$, the constraint $s_1 - s_2 \leq -4$ is subsumed by the constraint $s_1 - s_2 \leq -e_1; e_1 \in [3, 5]$. This holds true for every pair of comparable constraints $l_1$ and $l_2$, i.e. either $l_1 \Rightarrow l_2$ or $l_2 \Rightarrow l_1$. In other words, the domination relationship imposes a total order on each set of comparable constraints between two jobs. We use lexicographical ordering to break ties. It follows that in each equivalence class of the partition imposed by the comparability relationship there is a unique constraint that dominates all other constraints in that class [5].
Definition 10. The unique elements $\kappa_{ij}^{k} \in S_{ij}^{k}$, $k = 1, \ldots, 4$, which dominate all the other constraints in their respective partitions are called the dominators of that partition.

Lemma 1. If a constraint $l_1$ dominates another constraint $l_2$, then eliminating $l_2$ from the set of constraints, does not alter the parametric schedulability of the system i.e. if the constraint system has a parametric schedule with $l_2$ then it has a parametric schedule without $l_2$ and vice versa.

In other words, we can eliminate $l_2$ from the constraint set and test for parametric schedulability on the reduced set of constraints. In fact, it suffices to retain the 4 dominators between each pair of jobs, since all other constraints are redundant.

Proof: We provide a proof for the case in which both $l_1$ and $l_2$ belong to the set $S_{ij}^{1}$. The other three cases, viz. $l_1, l_2 \in S_{ij}^{2}, l_1, l_2 \in S_{ij}^{3},$ and $l_1, l_2 \in S_{ij}^{4}$ can be proved in identical fashion.

1. Let the initial system have a parametric schedule - We need to prove that the system will continue to be parametrically schedulable, even after $l_2$ is eliminated. But this is obvious, since we are reducing the number of constraints! (If $A_1$ is the feasible region, before the removal of $l_2$ and $A_2$ is the feasible region after the removal of $l_2$, then $A_1 \subseteq A_2$)

2. Let the initial system be parametrically unschedulable - We need to show that the removal of $l_2$ does not make the system schedulable. From our assumption $l_2 \in S_{ij}^{1}$, i.e. $l_2$ is a difference constraint that specifies decreased separation between the jobs $J_i$ and $J_j$. Since $l_1$ dominates $l_2$, the separation specified by $l_1$ is clearly smaller than the separation specified by $l_2$. Let us assume the contrary and suppose that the constraint system $A'$ resulting from the elimination of $l_2$ is parametrically schedulable. This means that there do not exist start times $s_i, s_j$ and execution times $e_i$ which could depend upon $s_i$ and $e_j$ which could depend upon both $s_i$ and $s_j$, such that a negative cost loop is created. (The system is infeasible if and only if there is such a negative cost loop \cite{15,17}.) Essentially, we are saying that $\forall s_i, e_i \in [l_i, u_i] \forall s_j, e_j \in [l_j, u_j] \ldots \neg((A, [s, e] \leq b) - \{l_2\})$ is false. This means that the constraints imposed by the sets $S_{ij}^{2}, S_{ij}^{4}$ can co-exist with $l_1$. Now consider what happens when $l_2$ is added to $A'$. $l_2$ cannot decrease the separation between jobs $J_i$ and $J_2$ any more than what is specified by $l_1$. Consequently if a negative loop could not be created using $l_1$, then such a loop definitely cannot be created using $l_2$, i.e. the constraint system stays parametrically schedulable, contradicting the assumption that it was not.

□

This leads us directly to:

Lemma 2. There are at most 4 non-redundant constraints between any 2 jobs; hence the total number of non-redundant constraints is at most $O(n^2)$. 
It follows that \texttt{Elim-Exist-Variable()} takes at most $O(n^2)$ time, since each start-time variable is part of at most $O(n)$ relationships. \texttt{Prune-Constraints()} basically performs the equivalent of finding the 4 maxima between each pair of start-times and hence the total time taken is proportional to the number of edges, which is $O(n^2)$. Checking the consistency of the resulting constraints can likewise be carried out in $O(n^2)$ time.

Since all the above functions are called at most $O(n)$ times, the above analysis leads to the following conclusion.

**Theorem 1.** Algorithm (5.1) can be implemented to run in $O(n^3)$ worst-case time.

**Proof:** Follows from the discussion above. $\square$

**References**


A Logical Framework for Knowledge Sharing in Multi-agent Systems

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Abstract. The issue of knowledge sharing has been an important topic in multi-agent research. Knowledge sharing leads to that agents analyze, judge and synthesize the told information so as to make agents’ own knowledge. To match these applications, this paper builds a logical framework for knowledge sharing among agents. We develop a multi-modal logic for reasoning about both agents’ knowledge and told information. For formalizing the relationship between knowledge and told information, we present a framework of semantics, with respect to which a sound and complete proof theory is given.

Keywords: logic, autonomous agents, philosophical foundations

1 Introduction

Currently, the research area of multi-agent systems is an active branch in artificial intelligence. In 1988, Bond and Gasser [1] defined a multi-agent system as a loosely coupled network of autonomous entities called agents which have individual capabilities, knowledge and resources, and which interact to share their knowledge and resources, and to solve problems being beyond their individual capabilities. From this definition, we can see that the issue of knowledge sharing is an important topic in multi-agent research. Therefore, many contributions have been made to this topic in the area of multi-agent systems, such as the knowledge sharing effort [14] and Knowledge Interchange Format (KIF) [15].

However, the issue of knowledge sharing among agents that usually have peculiarities of their knowledge and reasoning properties in the viewpoints of one agent has not been handled yet. We can use the following example to make clear what we exactly mean here. An agent $i$ may believe that another agent $j$’s knowledge or belief is veridical, i.e., in terms of standard multi-modal language, $K_i(K_j \alpha \Rightarrow \alpha)$, for all formulas $\alpha$. Only in this case, agent $i$ is willing to inherit all knowledge of agent $j$ and is ready to accept told information from agent $j$. So, it is important and interesting to present a reasoning model which enables agents to explicitly express, in their knowledge, both what they have been told and some peculiarities of their colleagues’ knowledge and reasoning properties. In this paper, we will tackle this issue.
On the other hand, although there has been a flourishing of epistemic logics formalizing multi-agent knowledge, belief, desire and intention, by using various kinds of possible-worlds semantics, despite their varieties of semantics, these logics usually assume a fixed epistemic logic for a certain class of modalities, and therefore all agents have the same reasoning properties. This is perhaps because the philosophical community, where various epistemic logics were firstly proposed, attacked and defended, has tended to concentrate on the single agent case. However, the case our formalism handles is about multiple heterogeneous agents.

In addition, the problem tackled in this paper is also interesting even in the case of fixing an epistemic logic or a class of epistemic properties. There is a few modal logics combining the notions of an agent’s belief (or knowledge), desire and intention; and there even exist general methods of fusing different modal logics (see [4]). However, so far there has not been yet one combining the notions of an agent’s knowledge and told information, which is thought surprisingly subtle whenever more than one agent is in the picture [5].

As a first step, we focus on the following epistemic properties:

- **Veridicality.** Only true things are known.
- **Introspection.** An agent is aware of what it knows and of what it does not know.
- **Consistency.** An agent’s knowledge is non-contradictory.

Note that, by considering different combinations of the above properties, the well-known modal logics \(S5, K, T, K45\) and \(KD45\), have been constructed and widely discussed.

In this paper, we aim to develop a model for reasoning about not only multi-agent knowledge and told information but also some key reasoning properties of agent’s knowledge, including the veridicality and the introspection. We allow the explicit mention of the introspection and the veridicality of an agent’s reasoning about its own knowledge, so that our language has formulas indicating which logic of \(K, T, S5, K45\) is used by an agent.

The rest of this paper is organized as follows. In next section, we review some basic concepts and notations of modal logic, which we will use through this paper. In Section 3, based on the analysis of relationship between knowledge and told information, we define the framework of semantics for our logic. Section 4 addresses the issue of proof theory of our logic. Finally, we discusses the importance of our logic framework in Section 5 and summarizes our contributions in Section 6.

### 2 Preliminaries

In this section, we recall some basic concepts and notations related to modal logics of multi-agents’ knowledge, which we shall use throughout the paper.

The language of those logics is propositional logic augmented by the modal operators \(K_1, \cdots, K_n\), where \(K_i\phi\) can be read “agent \(i\) knows (or believes) \(\phi\)".
We denote this language by $L_n$. For convenience, we define $true$ as an abbreviation for a fixed valid propositional formula, say $p \lor \neg p$, where $p$ is primitive proposition. We abbreviate $\neg true$ by $false$.

According to [5], semantics of these formulas can be given by means of Kripke structure [8], which formalize the intuition behind possible worlds. A Kripke structure is a tuple $(W, \pi, K_1, \cdots, K_n)$, where $W$ is a set of worlds, $\pi$ associates with each world a truth assignment to the primitive propositions, so that $\pi(w)(p) \in \{true, false\}$ for each world $w$ and primitive proposition $p$, and $K_1, \cdots, K_n$ are binary accessibility relations. By convention, $K_i^M$ and $\pi^M$ are used to refer to the $K_i$ relation and the $\pi$ function in the Kripke structure $M$, respectively. We omit the superscript $M$ if it is clear from context. Finally, we define

$$K_i(w) = \{w' | (w, w') \in K_i\}.$$ 

That is, $K_i(w)$ is the set of worlds agent $i$ considers possible in world $w$.

A situation is a pair $(M,w)$ consisting of a Kripke structure and a world $w$ in $M$. By using situations, we can inductively give semantics to formulas as follows: for primitive propositions $p$,

$$(M,w) \models p \text{ iff } \pi^M(w)(p) = true.$$ 

Conjunctions and negations are dealt with in the standard way. Finally,

$$(M,w) \models K_i \alpha \text{ iff for all } w' \in K_i^M(w), (M,w') \models \alpha.$$ 

Thus, an agent $i$ knows $\alpha$ if $\alpha$ is true in all situations that the agent considers possible.

Note that the Kripke structure $M$ is fixed in the above inductive interpretation. However, it is interesting that our interpretation, as given in the next section, is not in the case.

3 Non-standard Interpretation

This section presents the framework of semantics for our logic. We focus two classes of formulas, i.e., those of the form $K_i \phi$, and those of the form $T_i \phi$, which is read “agent $i$ is told $\phi$”. We use Kripke structures $(W, \pi, K_1, \cdots, K_n)$ and the formulas of the form $T_i \phi$ are interpreted by using the accessibility relation $K_i$ as usual, whereas the interpretation of the formulas of the form $K_i \phi$ is given by the analysis of relationship between knowledge and told information.

3.1 Language

The language we use, denoted by $L_n^T(VI)$, is $L_n$ augmented by one class of modalities $T_i$ and two classes of special proposition constants: $I_i$, and $V_i$ ($1 \leq i \leq n$), corresponding to the reasoning properties: the introspection and the veridicality, respectively. Thus, the formula $I_i$ says that agent $i$ is able to respect his knowledge; the formula $V_i$ indicates that the agent knows only true things.
We note that it is unnecessary for the language to include some other constants to express the consistency of agent $i$’s knowledge since we can use the formula $\neg K_i \text{false}$.

### 3.2 For Formulas of the Form $T_i \phi$

Now we indicate how to interpret formulas of $L_n(\text{VI})$. Let $\models_N$ be the satisfaction relation we are going to define. We here use standard Kripke structures $(W, \pi, K_1, \ldots, K_n)$ and situations. The key point is that the accessibility relation $K_i$ in Kripke structure is, however, no longer related to agent $i$’s knowledge; in a each situation $(M, w)$, the syntactic counterpart of those situations $(M, w')$, where $w' \in K_i(w)$, is what agent $i$ has been told rather than his knowledge.

Thus, given a situation $(M, w)$, we define

$$(M, w) \models_N T_i \phi \text{ iff for all } w' \in K_i(w)(M, w') \models_N \phi.$$ 

According to the above definition, each modality $T_i$ need not satisfy veridicality axioms and introspection axioms. Intuitively, the reason for non-introspection is clear; we can certainly tell or not tell you something without telling you that we have told or not told it. Assuming that the external state of agent $i$’s remains unchanged since the agent was told, it seems that the veridicality depends only on the honest of the teller. However, we argue that even if the teller is honest, the veridicality of modality $T_i$ is not necessarily true for $i$-subjective sentences since the subjective state of agent $i$ described by $i$-subjective sentences can be changed after agent $i$ is told something. For example, once agent $i$ is told “$p$ is true but you do not know it”, this sentence does not hold any longer, because agent $i$’s being told changes agent $i$’s subjective state.

### 3.3 For Formulas of the Form $K_i \phi$

The relationship between agents’ knowledge and the information that agents have been told has received a great deal of attention [11], and is thought of surprisingly subtle whenever more than one agent is considered in the picture [5]. Before interpreting the formulas of the form $K_i \phi$, we give some informal remarks on the relationship. Generally, the told information may be subject to noise; and thus, in order to obtain their own knowledge, agents need to refine the told information via various kinds of refining processes. Of these refining processes, we focus on the following three cases for the purpose of this paper:

- If agent $i$ is veridical, then the agent has the refining process of deleting the false pieces of the told information. Thus, even if agent $i$ is told some fact $p$, agent $i$ does not necessarily know $p$ whenever $p$ is false and agent $i$ is veridical.
- There are, for each agent, some special pieces of information, usually concerning the agent himself, that are straightforward accessible just by his/her feeling or observation rather than by his/her or other agents’ reasoning. Agents believe or refute such pieces of information firmly, no matter what has been told to them. For example, if you feel cold, then you know it, and
you still know it after you are told that you feel warm. The introspection property here is thought of as such information; if agent $i$ is able to introspect, then agent $i$ must know it, no matter what has been told to him.

- If agent $i$ is introspective, then the agent has the refining process of introspection. This process is surprisingly dependent on the veridicality of agent $i$, which leads the relationship between the told information and the agent’s knowledge somewhat subtle. In the absence of the veridicality, the introspection of agent $i$ means that if agent $i$ is (resp. is not) told something, then the agent must know he/she is (resp. is not) told it. In the presence of veridicality, agent $i$ pays his/her attention only to true things. The introspection of agent $i$ thus indicates that (a) if agent $i$ is told a proposition $p$ which is true, then agent $i$ knows he/she is told the fact, and (b) if a proposition $p$ is false, then agent $i$ knows that he/she is not told $p$ even though he is indeed told $p$. Nevertheless, in both the cases, we have that if agent $i$ knows something, then he/she knows he/she knows it, and if he/she does not know it, then he/she knows he/she does not. Therefore, the meaning of introspection we use here is consistent with that in the literature.

Now let us consider the interpretation of the formulas of form $K_i \alpha$. It is crucial to ask what situations agent $i$ should think possible from the standpoint of what he/she actually knows after checking and cogitating on the told information. Let $S_i(M, w)$ be the set of all such possible situations for the given situation $(M, w)$. Then, we interpret formulas of the form $K_i \alpha$ in this way:

$$(M, w) \models_N K_i \alpha \iff \text{for all } (M', w') \in S_i(M, w), (M', w') \models_N \alpha.$$ 

We now figure out what $S_i(M, w)$ should be. There are four cases according to the logical style of agent $i$ in the situation $(M, w)$. Firstly, if the logical style is $K$, i.e. the value of $\pi(w)$ is false at both $I_i$ and $V_i$, then agent $i$ is unable to distinguish his/her knowledge from what has been told, and hence the situations in $S_i(M, w)$ are exactly those $(M, w')$’s, where $w' \in K_i(w)$.

Secondly, assume the logical style is $T$, i.e.

$$\pi(w)(I_i) = \text{false} \text{ and } \pi(w)(V_i) = \text{true},$$

then, we get $S_i(M, w)$ by adding the actual world $w$ to each $(M, w')$, where $w' \in K_i(w)$. This enables agent $i$ to delete the false pieces of the told information.

Thirdly, suppose the logical style is $K45$, i.e.

$$\pi(w)(I_i) = \text{true} \text{ and } \pi(w)(V_i) = \text{false}.$$ 

Assuming agent $i$ thinks situations $(M', w')$ possible from his own knowledge, we have, by the introspection of agent $i$, that agent $i$’s told information in each of those situations $(M', w')$ is exactly that in the actual situation $(M, w)$, which is semantically represented as

$$K_i^{M'}(w') = K_i^M(w).$$
On the other hand, as we mentioned above, the introspection property \( I_i \) is such information that if it holds, then agent \( i \) must know it, no matter what agent \( i \) has been told. Thus, in each of those situations \((M', w')\), \( I_i \) must hold, hence

\[
\pi^{M'}(w')(I_i) = \text{true}.
\]

Based on the above discussions, we define \( S_i(M, w) \) as the set of those situations \((M', w')\), where \( w' \in \mathcal{K}^M_i(w) \) and \( M' \) coincides with \( M \) but

\[
\mathcal{K}^{M'}_i(w') = \mathcal{K}^M_i(w) \text{ and } \pi^{M'}(w')(I_i) = \text{true}.
\]

For convenience, we denote the above \( M' \) Kripke structures by

\[
M \left[ \frac{\mathcal{K}_i(w')}{\mathcal{K}_i(w)} \right] \frac{\pi(w')(I_i)}{\text{true}}.
\]

Finally, let the logical style be \( \mathbf{S5} \). Then, by considering the veridicality, we should put the actual world \( w \) into the set \( \mathcal{K}^M_i(w) \), and think that \( \mathcal{K}^M_i(w) \cup \{w\} \) were the set of worlds possible from the standpoint of agent \( i \)'s told information. By considering the introspection, we define \( S_i(M, w) \) in the same way as in the case of \( \mathbf{K45} \), nevertheless \( \mathcal{K}^M_i(w) \), the set of worlds possible from the standpoint of agent \( i \)'s told information, is replaced by \( \mathcal{K}^M_i(w) \cup \{w\} \). In other words, we define \( S_i(M, w) \) as the set of those situations \((M', w')\), where \( w' \in \mathcal{K}^M_i(w) \cup \{w\} \), and \( M' \) coincides with \( M \) but

\[
\mathcal{K}^{M'}_i(w') = \mathcal{K}^M_i(w) \cup \{w\} \text{ and } \pi^{M'}(w')(I_i) = \text{true}.
\]

For convenience, we denote the \( M' \) Kripke structures by

\[
M \left[ \frac{\mathcal{K}_i(w')}{\mathcal{K}_i(w) \cup \{w\}} \right] \frac{\pi(w')(I_i)}{\text{true}}.
\]

To summarize,

\[
S_i(M, w) = \begin{cases} 
\{(M, w') \mid w' \in \mathcal{K}^M_i(w)\}, & \text{if } \pi(w)(I_i) = \text{false} \text{ and } \pi(w)(V_i) = \text{false}; \\
\{(M, w') \mid w' \in \mathcal{K}^M_i(w) \cup \{(M, w)\}\}, & \text{if } \pi(w)(I_i) = \text{false} \text{ and } \pi(w)(V_i) = \text{true}; \\
\{(M \left[ \frac{\mathcal{K}_i(w')}{\mathcal{K}_i(w)} \right], \pi(w')(I_i)), w' \mid w' \in \mathcal{K}^M_i(w)\}, & \text{if } \pi(w)(I_i) = \text{true} \text{ and } \pi(w)(V_i) = \text{false}; \\
\{(M \left[ \frac{\mathcal{K}_i(w')}{\mathcal{K}_i(w) \cup \{w\}} \right], \pi(w')(I_i)), w' \mid w' \in \mathcal{K}^M_i(w) \cup \{w\}\}, & \text{if } \pi(w)(I_i) = \text{true} \text{ and } \pi(w)(V_i) = \text{true}. 
\end{cases}
\]

### 3.4 Satisfaction Relation

Now we formally present our semantics framework by defining inductively the satisfaction relation \( \models_N \) between a situation and a formula as follows. Firstly,
(M, w) \models_N p \text{ iff } \pi(w)(p) = \text{true}, for primitive propositions p. Secondly, 
(M, w) \models_N \neg \alpha \text{ iff not } (M, w) \models_N \alpha, and (M, w) \models_N \alpha \land \beta \text{ iff } (M, w) \models_N \alpha \text{ and } (M, w) \models_N \beta. Finally, (M, w) \models_N K_i \alpha \text{ iff } (M', w') \models_N \alpha \text{ for all } (M', w') \in S_i(M, w).

We remark that according to this semantics, a veridical agent takes the information he/she was told and adds the current world to the set of possibilities to obtain its knowledge. However, it seems difficult to construct such agents because an agent typically does not know which is the current world and it is impossible for an agent to perform this operation of adding the current world. Nevertheless, this does not imply this semantics is counter-intuitive and susceptible. There might be no veridical agent in the real world, but there are agents that do have the veridicality properties from some other agents’ viewpoints. For example, if agent j thinks that agent i is veridical, then, from agent j’s viewpoint, agent i does not yet know the current world, but agent j knows (or believes) that the current world is one of agent j’s possible worlds. Thus, if the agent j supposes that the current world is w and agent i’s set of possible worlds (corresponding to its told information) is W, then agent j thinks that agent i’s knowledge is determined by the set W adding the supposed current world w.

4 Proof Theory

This section gives the proof theory of our logic. Let $K_n(T)$ be the well-known system $K_n$ for the modalities $T_i$ instead of $K_i$. Clearly, the axioms $P$ and $K$ and the inference rules in the system $K_n(T)$ are valid with respect to the semantics presented above. In addition, in the absence of $K_i$, the system $K_n(T)$ is also complete with respect this semantics. Furthermore, we capture the following important properties of the relationship between told information and knowledge.

**Theorem 1** Let $\phi$ and $\varphi$ be arbitrary formulas of $L^T_n(VI)$, and $\psi$ $i$-objective. Then the following formulas are valid, i.e., satisfied in all situations $(M, w)$:

1. $(K_i \phi \land K_i(\phi \Rightarrow \varphi)) \Rightarrow K_i \varphi$
2. $(\neg I_i \land \neg V_i) \Rightarrow (K_i \phi \Leftrightarrow T_i \phi)$
3. $(\neg I_i \land V_i) \Rightarrow (K_i \phi \Leftrightarrow (T_i \phi \land \phi))$
4. $I_i \Rightarrow K_i I_i$
5. $(I_i \land \neg V_i) \Rightarrow (K_i \psi \Leftrightarrow T_i \psi)$
6. $(I_i \land V_i) \Rightarrow (K_i \psi \Leftrightarrow T_i \psi \land \psi)$
7. $(I_i \land \neg V_i) \Rightarrow ((T_i \phi \Rightarrow K_i T_i \phi) \land (\neg T_i \phi \Rightarrow K_i \neg T_i \phi))$
8. $(I_i \land V_i) \Rightarrow (((T_i \phi \land \phi) \Rightarrow K_i T_i \phi) \land ((\neg T_i \phi \lor \neg \phi) \Rightarrow K_i \neg T_i \phi)))$

This theorem is easy to check by using routine method; we omit its proof here. Let $TK^VI_n$ be the system that results from adding the above valid formulas for each $i$ to the axioms for the system $K_n(T)$. Thus, Theorem [1] implies that $TK^VI_n$ is sound with respect to our semantics.

---

[1] We define $i$-objective formulas as Boolean combinations of primitive propositions (except $I_i$) and formulas of the form $T_j \phi$, $j \neq i$. 

Theorem 2  For each formula $\phi$ in $L^T_n$(VI), there is a formula $\phi'$ such that in the system $TK^V^n$, the two formulas are provably equivalent, but the latter does not have any appearance of $K_i$'s.

The detail of this proof is omitted for limited space.

For convenience, for an arbitrary formula $\varphi$ in $L^T_n$(VI), by $\overline{\varphi}$ we denote a formula which provably equivalent to $\varphi$ and has no appearance of $K_i$'s. By the above result it follows immediately the following theorem.

Theorem 3 (Soundness and completeness) For the formulas in $L^T_n$(VI), the system $TK^V^n$ is a sound and complete axiomatization with respect to the semantics above.

Proof. The proof of the soundness is easy; axioms $TK_1 \cdots TK_8$ are valid by Theorem 1 and other axioms and inference rules are trivially true. As for the completeness, assume that $\Gamma$ is an arbitrary set of consistent formulas. We must prove that $\Gamma$ is satisfiable by some Kripke structure. Let the formula set $\Gamma' = \{ \overline{\varphi} \mid \varphi \in \Gamma \}$. By Theorem 2, $\Gamma$ is provably equivalent to $\Gamma'$, and it suffices to show $\Gamma'$ is satisfiable by some Kripke structure. But it is well-known that $K_n(T)$ is complete with respect to general Kripke structures; we therefore conclude that $\Gamma'$ is satisfiable by a Kripke structure. □

We end this section by pointing out that

$$(I_i \land V_i) \Rightarrow (K_i \neg T_i \phi \iff (\neg T_i \phi \lor \neg \phi))$$

is valid, and thus the formula $K_i \neg T_i p$ does not necessarily imply $\neg T_i p$ even under the condition $(I_i \land V_i)$. It follows that Axiom $T$ doesn’t hold for our extended language, which makes this system novel in contrast to usual combinations of two classes of modalities (or modal logics).

This also sheds some light on the appropriateness of $S5$ as a logic of knowledge. The introspection axiom of $S5$ has been attacked at some length in the literature of philosophy and artificial intelligence (see, for example, [10]). Our research indicates that

- for veridical and introspective agents, logic $S5_n$ is an appropriate logic of knowledge; and
- the introspection axiom reasonably stands even if there are other interactive modalities in the picture.

Nevertheless, the veridicality axiom, which was not attacked before, does not necessarily hold in the presence of other modalities with respect to our semantics.

5  Discussions

This section further shows the significance of our logic in the both philosophical and AI respects.
5.1 General Logic of Knowledge

Consider the language $\mathcal{L}_n^T(VI)$ that consist of those formulas in $\mathcal{L}_n^T(VI)$ without any appearance of modalities $T_i$. Our semantics can be used to interpret those formulas. This yields a general logic of knowledge such that different agents may have different reasoning styles.

The approach to the above purpose seems to be straightforward via multiple versions of Kripke structures in which different agents can have different types of accessibility relations. However, in such an approach, the reasoning style of a fixed agent, from viewpoints of different agents, remains unchanged. Actually, it is viewed as common knowledge among all agents concerned. This leads to simple classification of agents provided by three properties: veridicality, introspection and consistency. Such a classification is not appropriate for most general systems of multiple agents. For example, very few agents are of veridicality property in every context.

5.2 Only Knowing

The notion of only knowing (see [12]) plays an important role in knowledge representation, which closely related to formalisms of nonmonotonic reasoning, such as default logic [16] and circumscription [13]. Several notions of multi-agent only knowing have been developed under the assumption that all agents use the same logic [9,6,7]. However, in the context of heterogeneous systems, this assumption is no longer appropriate.

Based on the results in this paper, we will propose a general notion of multi-agent only knowing. In our approach, agent $i$ only knowing something results from the agent only being told something, which is easier to formalize because the logic styles for the modality $T_i$'s are fixed $K$ despite the multi-formity of logics of agents’ knowledge. This is also intuitively reasonable; after all, the less is told, the less is known.

6 Conclusion

In this paper, we have developed a model for reasoning about not only agents’ knowledge and told information but also some key reasoning properties of agent’s knowledge such as the introspection. These reasoning properties are regarded as styles of how an agent transforms the rough information that has been told, into its own knowledge after checking and cogitating on the former. By using Kripke structures abnormally, we have presented a new framework of semantics, and we give a sound and complete proof theory with respect to the semantics.

Future work includes building more realistic models of knowledge and communication, where more parameters of properties of accessibility relations may be taken into account. It is also interesting to consider temporal factors in those models. Moreover, it is worth to explore the notion of only knowing derived by that of only being told, and compare them with those of previous treatments [12,9,6] on the notion of only knowing.
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References
A Lockout Avoidance Algorithm without Using Time-Stamps for the $k$-Exclusion Problem

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Abstract. We propose a lockout avoidance algorithm for the $k$-exclusion problem on the asynchronous multi-writer/reader shared memory model. This algorithm is the extension of an accelerated version of the $n$-process algorithm to the $k$-exclusion problem. The running time by the algorithm for the trying region of any faultless process is bounded by $(n - k)c + O(n(n - k)^2)l$ if the number of stopping process failures is less than $k$, where $n$ is the number of processes, $l$ is an upper bound on the time between two successive atomic steps for any faultless process, and $c$ is an upper bound on the time that any user spends in the critical region.

key words: asynchronous processes, concurrent computation, $k$-exclusion, lockout avoidance, lockout freedom, shared memory

1 Introduction

Mutual exclusion is a problem of managing access to a single indivisible resource that can only support one user at a time. The $k$-exclusion problem is a natural extension of the mutual exclusion problem. In $k$-exclusion, some number of processes, specified by the parameter $k$, are allowed to be concurrently inside the critical region, where corresponding users can use the resource. This extension was first defined and solved by Fischer, et al. on the shared memory model [4]. Afek, et al. [1] gave another solution to the $k$-exclusion problem. Unlike the solution by Fischer et al. this solution does not use a powerful read-modified-write primitive. By the use of a concurrent time-stamp system [2,3], it requires only bounded read/write shared memory [1].

The $n$-process algorithm for the mutual exclusion problem was proposed by Peterson [7], and accelerated by Igarashi and Nishitani [5]. The algorithm proposed in this paper is an extension of these algorithms to the $k$-exclusion problem on the asynchronous multi-writer/reader shared memory model.

The algorithm is immune to stopping failures of fewer than $k$ processes. All logarithms used in this paper are to the base 2. We impose an upper bound $l$ on the time between successive atomic steps of each faultless process in the trying region and the exit region, and an upper bound $c$ on the time that any user spends in its critical region. Let $n$ be the number of processes in a system. The running time by the algorithm is bounded by $(n - k)c + O(n(n - k)^2)l$. When $k = 1$ (i.e., the mutual exclusion case), the running time $(n - 1)c + O(n^3)l$ of the
algorithm proposed in this paper is better than the running time $O(n^2)c + O(n^4)l$
of the $n$-process algorithm by Peterson [5,7] and as fast as the running time of
the accelerated versions by Igarashi and Nishitani [3].

2 Preliminaries

The asynchronous multi-writer/reader shared memory model consists of pro-
cesses and shared variables. Interactions between a process and its correspond-
ing user are by input actions from the user to the process and by output actions
from the process to the user. All communication among the processes is via the
shared memory. The model is called an I/O automaton [6].

A user with access to the resource is modeled as being in a critical region.
When a user is not involved in the resource, it is said to be in its remainder
region. In order to gain admittance to its critical region, a process executes a
trying protocol. The duration from the start of execution of the trying protocol
to the entrance of the critical region is called the trying region. After the end
of the use of the resource by a user, the corresponding process executes an exit
protocol. The duration of the execution of the exit protocol is called the exit
region. The $k$-exclusion problem is to devise protocols for at most $k$ processes
to be concurrently in the critical regions.

We assume that the $n$ processes are numbered $1, \ldots, n$. Each process $i$
corresponds to user $U_i$ ($1 \leq i \leq n$). The inputs to process $i$ from user $U_i$ are try$_i$
which means a request by user $U_i$ for access to the resource, and exit$_i$ which
means an announcement of the end of the use of the resource by $U_i$. The output
from process $i$ to user $U_i$ is crit$_i$ which means the grant of the resource to $U_i$, and
rem$_i$ which tells $U_i$ that it can continue with the rest of its work.

The following conditions should be satisfied.

(1) There is no reachable system state in which more than $k$ processes are in
the critical regions.
(2) If at least one faultless process is in the trying region and less than $k$
processes are in the critical region, then at some later point some process enters its
critical region.
(3) If a faultless process is in the exit region, then at some later point the process
enters its remainder region.
(4) If all users always return the resource and if the number of faulty processes
of the stopping type is at most $f < k$, then any faultless process wishing to
enter the critical region eventually does so.

Conditions (1), (2), (3) and (4) above are called respectively $k$-exclusion,
progress for the trying region, progress for the exit region, and $k$-lockout avoid-
ance.
3 A Lockout Avoidance Algorithm

procedure \((n,k)\)-EXCL

shared variables

for every \(s \in \{1, \ldots, n-k\}\):

\(\text{turn}(s) \in \{1, \ldots, n\}\), initially arbitrary, writable and readable by all processes;

for every \(i \in \{1, \ldots, n\}\):

\(\text{flag}(i) \in \{0, \ldots, n-k\}\), initially 0, writable by \(i\) and readable by all \(j \neq i\);

process \(i\)

input actions \{inputs to process \(i\) from user \(U_i\)\}:

\(\text{try}_i, \text{exit}_i\);

output actions \{outputs from process \(i\) to user \(U_i\)\}:

\(\text{crit}_i, \text{rem}_i\);

** Remainder region **

\(\text{try}_i;\)

for \(s = 1\) to \(n-k\) do

begin

\(\text{flag}(i) := s;\)

\(\text{turn}(s) := i;\)

waitfor \(\left|\{j|j \neq i : \text{flag}(j) \geq s\}\right| \leq n - s - 1\) or \(\text{turn}(s) \neq i\)

end;

\(\text{crit}_i;\)

** Critical region **

\(\text{exit}_i;\)

\(\text{flag}(i) := 0;\)

\(\text{rem}_i;\)

For each level \(s\) \((1 \leq s \leq n-k)\), the statement waitfor\(\left|\{j|j \neq i : \text{flag}(j) \geq s\}\right| \leq n - s - 1\) or \(\text{turn}(s) \neq i\) in \((n,k)\)-EXCL is not an atomic step. It consists of \(O(n)\) atomic steps. For each \(s\), the contents of the local variable \(\text{count}_i\) does not necessarily represent the number of other processes that are located at level \(s\) or above. Since the shared memory used in this paper is not the read-modified-write model, this uncertainty cannot be avoided.

4 Correctness of \((n,k)\)-EXCL

We may assume that any process is initially in the remainder region. Throughout this paper, we assume that the number of stopping process failures is always at least \(k - 1\). In an execution by \((n,k)\)-EXCL, process \(i\) is said to be a winner at level \(s\) if it has left the waitfor statement in the \(s\)th loop of the for statement. Note that if a process is a winner at level \(s\) then for any \(1 \leq t \leq s\), the process is also a winner at level \(t\). For each \(i\) \((1 \leq i \leq n)\), when process \(i\) has entered its exit region, the qualification for the winner of process \(i\) is canceled. We omit the proof of the following lemma due to the space limitations.
Lemma 1. In any reachable system state of \((n, k)\)-EXCL, for any \(1 \leq s \leq n-k\) there are at most \(n-s\) winners at level \(s\).

The next theorem follows immediately from Lemma 1.

Theorem 1. \((n,k)\)-EXCL guarantees \(k\)-exclusion even if any number of process failures of the stopping type exist.

It is obvious that \((n, k)\)-EXCL guarantees the progress for the exit region. In order to prove the progress for the trying region, and \(k\)-lockout avoidance, it is enough to give a time bound for the trying region.

Theorem 2. Suppose that the number of stopping failures of processes is always at most \(k-1\). In \((n, k)\)-EXCL, the time from when a faultless process enters its trying region until the process enters its critical region is bounded by \((n-k)c + O(n(n-k)^2)l\).

Proof. For each \(s\), \(1 \leq s \leq n-k\), define \(F(s)\) to be the maximum time from when a faultless process has entered the competition at level \(s\) until it becomes a winner at level \(s\). The worst situation for a faultless process at level \(n-k\) to reach the critical region is the following case. Consider the case where there are \(k+1\) winners at level \(n-k-1\) and all of them are competitors including the process itself at level \(n-k\), and then the process becomes a loser among them at level \(n-k\).

In the worst case, the time to decide \(k\) winners at level \(n-k\) is at most \((2n+3)l\). The time duration includes 2 atomic steps for setting flag values and setting turn\((n-k)\) at level \(n-k\), the time to test the condition in the waitfor statement at level \(n-k\), and local computing time for each faultless process to count the number of flag\((j)\)'s with values not less than \(n-k\) and for checking the value of turn\((n-k)\). Since the number of stopping failures of processes is at most \(k-1\), it is not possible that all the \(k\) winners are faulty processes. The time from the end of the competition at level \(n-k\) until the time when at least one faultless winner resets its flag value to 0 in its exit region is at most \(3l + c\). Then the loser at level \(n-k\) becomes a winner at the level within \((2n+1)l\) time after the reset of the winner’s flag value. Hence, \(F(n-k) \leq (4n + 7)l + c\).

The worst situation for a faultless process at level \(s\), \(1 \leq s \leq n-k-1\), to become a winner at level \(s\) is the case where \(n-s+1\) processes have just entered the competition at level \(s\), or some of the processes have stopped due to their stopping failures after they had become winners at level \(s\), whereby the faultless process becomes a loser at level \(s\). In this case, the time required to decide the winners at level \(s\) is also at most \((2n+3)l\). The faultless loser at level \(s\) will become a winner when a new process joins the competition at level \(s\) and then sets turn\((s)\) to be its process name, or when at least one winner at the level eventually resets its flag value to 0 in its exit region. In the worst case, it is sufficient to consider the latter case. In order that the loser at level \(s\) becomes a winner at that level, it must wait until at least one winner at level \(s\) exits from the critical region. Since the number of faulty processes is at most \(k-1\), there exists a faultless winner at that level. The quickest winner at level \(s\) can reach the critical region within \((n-k-s)(2n+3)l\) time. Hence,
the loser at level \( s \) can move up as a winner at level \( s + 1 \) within
\[
F(s) \leq (n - k - s + 1)(2n + 3)l + (2n + 1)l + 3l + c = (n - k - s + 2)(2n + 3)l + l + c.
\]
Thus the time from when a faultless process has entered its trying region until it enters its critical region is bounded by
\[
\sum_{s=1}^{n-k} F(s) \leq \sum_{s=1}^{n-k-1} [(n - k - s + 2)(2n + 3)l + l + c] + (4n + 7)l + c \\
\leq (n - k)c + O(n(n - k)^2)l.
\]

The next theorem follows immediately from Theorem 2 and Theorem 3.

**Theorem 3.** \((n, k)\)-EXCL solves the \( k \)-exclusion problem, and it is \( k \)-lockout avoidance.

5 Concluding Remarks

The total size of the shared memory for \((n, k)\)-EXCL is \((n - k)\lceil \log n \rceil + n\lceil \log(n - k + 1) \rceil \) bits. We are interested in a problem of whether some modification of these algorithms can reduce the shared memory size. We are also interested in a problem of whether we can speed up our algorithm by simple modification for the case where both \( k \) and \( n - k \) are \( \Theta(n) \) or for the case where \( k \) or \( n - k \) is \( O(1) \). The \( k \)-assignment problem is closely related to the \( k \)-exclusion problem. We are tempted to solve the \( k \)-assignment problem without using concurrent time-stamps by similar techniques.

References

Prefix-Free Languages and Initial Segments of Computably Enumerable Degrees

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Abstract. We study prefix-free presentations of computably enumerable reals. In [2], Calude et. al. proved that a real \( \alpha \) is c.e. if and only if there is an infinite, computably enumerable prefix-free set \( V \) such that \( \alpha = \sum_{\sigma \in V} 2^{-|\sigma|} \). Following Downey and LaForte [5], we call \( V \) a prefix-free presentation of \( \alpha \). Each computably enumerable real has a computable presentation. Say that a c.e. real \( \alpha \) is simple if each presentation of \( \alpha \) is computable. Downey and LaForte [5] proved that simple reals locate on every jump class. In this paper, we prove that there is a non-computable c.e. degree bounding no noncomputable simple reals. Thus, simple reals are not dense in the structure of computably enumerable degrees.

1 Introduction

This paper investigates the relationship between the c.e. degrees (the computational complexity classes under Turing reducibility of sets of positive integers which can be effectively enumerated) and the presentations of computably enumerable reals. Investigation on effectiveness of reals originated in Turing’s paper [11], and has now become one of the central topics of computable analysis and algorithmic information theory.

Say that a real \( \alpha \) is computable if \( \alpha \) is the limit of a computable sequence of rationals \( \{ r_i : i \in \omega \} \) and there is a computable algorithm \( M \) such that for all \( n \), \( |\alpha - r_M(n)| < 2^{-n} \). All algebraic numbers, \( \pi \), the Euler number \( e \), and all numbers commonly used in numerical analysis are computable. Note that \( M \) here controls the rate of the convergence of \( \{ r_i : i \in \omega \} \). Effectivity of \( M \) is crucial because there are computable sequences of rationals converging (non-effectively) to non-computable reals, see Spector [10]. Based on this, Soare [8] defined computably enumerable reals as follows:

Definition 1 We call a real \( \alpha \) computably enumerable (c.e., for short), if \( \alpha \) is the limit of a computable sequence of rationals \( \{ r_i : i \in \omega \} \).

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For simplicity, we only consider reals between 0 and 1. In [8], Soare proved that \( \alpha \) is c.e. iff the lower (left) Dedekind cut of \( \alpha \) is computably enumerable. For more equivalent definitions of c.e. reals, we need a technical notion whose use is crucial in the investigations of randomness of reals.

Let \( \Sigma = \{0, 1\} \) be the binary alphabet, and \( \Sigma^* \) be the set of finite binary strings. For \( \sigma \in \Sigma^* \), let \( |\sigma| \) be the length of \( \sigma \). A \( \subseteq \Sigma^* \) is prefix-free if for any \( \sigma_1, \sigma_2 \in A \), if \( \sigma_1 \) is a prefix of \( \sigma_2 \), then \( \sigma_1 = \sigma_2 \). Say that \( C \) is a Chaitin machine (self-delimiting Turing machine) if \( C \) is a Turing machine processing binary strings such that its program set (domain) \( \text{PROG}_C = \{ \sigma \in \Sigma^* | C(\sigma) \text{ halts} \} \) is prefix-free. Clearly, \( \text{PROG}_C \) is computably enumerable (c.e. for short). Conversely, each prefix-free c.e. set of strings is the domain of some Chaitin machine. The program-size complexity of the string \( \sigma \in \Sigma^* \) (relatively to \( C \)) is defined as \( H_C(\sigma) = \min\{ |\eta| : \eta \in \Sigma^*, C(\eta) = \sigma \} \), where \( \min \emptyset = \infty \).

**Theorem 1** (Invariance Theorem) There is a Chaitin machine \( U \) (called universal) such that for every Chaitin machine \( C \), \( H_U(\sigma) \leq H_C(\sigma) + O(1) \).

Note that \( \text{PROG}_U \) is c.e., but not computable.

Generally, for any prefix-free languages, \( A \) say, \( \sum_{\sigma \in A} 2^{-|\sigma|} \) converges to a real in \([0, 1]\), i.e., \( \sum_{\sigma \in A} 2^{-|\sigma|} \leq 1 \). This is called the Kraft’s Inequality. We use \( 2^{-A} \) to denote \( \sum_{\sigma \in A} 2^{-|\sigma|} \). Thus, by the property of \( \text{PROG}_U \) mentioned above, we have that \( 2^{-\text{PROG}_U} \) is a c.e., but noncomputable, real in the unit interval. As usual, we use \( \Omega_U \) to denote the real \( 2^{-\text{PROG}_U} \). Calude, Hertling, Khoussainov and Wang [2], Downey, Hirschfeldt and Nies [4], Kučera and Slaman [6] provide extensive investigations of the randomness of \( \Omega_U \).

The following are equivalent definitions of c.e. reals.

**Theorem 2** (Calude et. al. [2]) For any \( \alpha \) in the unit interval, the following are equivalent:

1. \( \alpha \) is computably enumerable.
2. The left Dedekind cut of \( \alpha \) is computably enumerable.
3. There is an infinite computably enumerable prefix-free set \( V \subset \Sigma^* \) such that \( \alpha = 2^{-V} \).
4. There is a computable prefix-free set \( V \subset \Sigma^* \) such that \( \alpha = 2^{-V} \).
5. There is a computable increasing sequence of rationals with limit \( \alpha \).

Based on characterization (3), Downey and LaForte [5] proposed the notion of presentations as follows:

**Definition 2** (Downey and LaForte [5]) Say that \( V \subseteq \Sigma^* \) is a prefix-free presentation of c.e. real \( \alpha \) if (1) \( V \) is a prefix-free c.e. set; (2) \( \alpha = 2^{-V} \).

By (4) of Theorem 2, each c.e. real has a computable presentation. Say that a c.e. real \( \alpha \) is simple if all presentations of \( \alpha \) are computable. Since simple c.e. reals resemble computable reals, it is tempting to conjecture that simple reals might be related to the known classes of c.e. sets resembling computable sets (like the low sets) in the sense of the jump hierarchy. Downey and LaForte [5] refuted this by showing that there are simple c.e. reals which are high.

**Theorem 3** (Downey and LaForte [5]) There is a simple real \( \alpha \) with \( \alpha' \equiv_T \emptyset'' \).
From the point of view of computability theory, Downey and LaForte’s result suggests to asking whether all c.e. degrees contain some simple reals, or whether the degrees of simple reals are dense in the class of all c.e. degrees. In this paper, we prove that answers to these two questions are both negative.

**Theorem 4** There is a noncomputable c.e. set $A$ such that $A$ bounds no noncomputable simple reals. That is, for any noncomputable c.e. real $\alpha \in [0,1]$, if $\alpha \leq_T A$, then $\alpha$ has a noncomputable presentation.

Here, $\alpha \leq_T A$ means that there is a (partial) computable functional $\Phi$ such that $\alpha = \Phi(A)$, namely, the real $0.\Phi(A;0)\Phi(A;1)\cdots$ is a binary expansion of $\alpha$. Theorem 4 is proved by induction as follows:

**Theorem 5** There is a noncomputable c.e. set $A$ s.t. for any $n \geq 0$, for any noncomputable c.e. real $\alpha \in [1-\frac{1}{2^n},1-\frac{1}{2^{n+1}}]$, if $\alpha \leq_T A$, then $\alpha$ has a noncomputable presentation.

The main part of this paper is devoted to the proof of the base step $n=0$:

**Theorem 6** There is a noncomputable c.e. set $A$ s.t. for any noncomputable c.e. real $\alpha \in [0,\frac{1}{2}]$, if $\alpha \leq_T A$, then $\alpha$ has a noncomputable presentation.

The successor step of the proof is: Fix $A$ as in Theorem 5, assume that we have Theorem 5 for $n$. We now consider the case $n+1$. Fix $\alpha \in [1-\frac{1}{2^n},1-\frac{1}{2^{n+1}}]$ and $\beta \leq_T A$ as a noncomputable c.e. real. Then $\beta = 2\alpha - 1 \in [1-\frac{1}{2^n},1-\frac{1}{2^{n+1}}]$ and $\beta \leq_T A$ is also a noncomputable c.e. real. By the induction hypothesis, $\beta$ has a noncomputable presentation $V_\beta$. Define $V_\alpha = \{0\sigma : \sigma \in V_\beta\} \cup \{1\}$. Then $V_\alpha$ is c.e. and noncomputable. Prefix-freeness of $V_\alpha$ comes from that of $V_\beta$ directly. Also we have

$$2^{-V_\alpha} = \sum_{\eta \in V_\alpha} 2^{-|\eta|} = \sum_{\sigma \in V_\beta} 2^{-|0\sigma|} + \frac{1}{2} = \frac{1}{2} \sum_{\sigma \in V_\beta} 2^{-|\sigma|} + \frac{1}{2} = \frac{1}{2} \beta + \frac{1}{2} = \alpha.$$ 

Therefore, $V_\alpha$ is a presentation of $\alpha$. \qed

## 2 Requirements

In this section we list the requirements necessary to prove Theorem 6. We will construct a c.e. set $A$ satisfying for all $i$ the requirements:

- $P_i$: $A \neq W_i$;
- $R_i$: $[\alpha_i = \Phi_i(A) = 2^{-F_i} \in [0,\frac{1}{2}]$ & $\alpha_i$ is c.e. & $F_i$ is prefix-free] $\Rightarrow$ $[\alpha_i$ is computable $\lor (\exists$ noncomputable c.e. $U_i)(U_i$ presents $\alpha_i)]$.

Here we assume that $\{W_i\}_{i \in \omega}$ is a listing of all c.e. sets of natural numbers, and $\{(\alpha_i,\Phi_i,F_i)\}_{i \in \omega}$ is a listing of all triples, $(\alpha,\Phi,F)$ such that $\alpha$ is a c.e. real, $\Phi$ is a partial computable functional, and $F$ is a c.e. set of strings. Since $R_i$ is a complicated requirement, we break it into simpler subrequirements $S_{i,j}$ as follows:

- $S_{i,j}$: $U_i \neq V_j$.

Here $\{V_i\}_{i \in \omega}$ is a listing of all c.e. sets of strings in $\Sigma^*$. 
3 Basic Modules for Each Requirement

We now describe the atomic strategies (modules) for each type of requirement $P_i, R_i, S_{i,j}$.

(3.1) The $\rho$-module for $P_i$

A $P$-strategy, $\rho$ say, trying to figure out a number $x$ such that $A(x) \neq W_i(x)$, is the standard Friedberg-Muchnik procedure. $\rho$ runs as follows:

1. Choose a big number, $x$ say.
2. Wait for $\Phi_i(x) \downarrow$.
3. Put $x$ into $A$, and stop.

$\rho$ has two finitary outcomes: $w$ and $s$, where $w$ denotes the fact that $\rho$ waits at (2) forever, and $s$ denotes the fact that $\rho$ reaches (3) and stop. In both cases, $A(x) \neq W_i(x)$, and $P_i$ is satisfied.

(3.2) The $\tau$-module for $R_i$

Let $\alpha_i[s]$ be the approximation of $\alpha_i$ at stage $s$, $F_i$ is a c.e. set of strings presenting $\alpha_i$. W.o.l.g., we assume that for all $s$, $\alpha_i[s] \leq \sum_{\sigma \in F_i,s} 2^{-|\sigma|} \leq \frac{1}{2}$.

Define the distance between $\alpha_i$ and $\Phi_i^A$ as follows:

$$\ell(i, s) = \max\{x < s : (\forall y < x)[\Phi_i^A(y)[s] \downarrow]\};$$
$$m\ell(i, s) = \max\{0, \ell(i, t) : t < s\},$$
$$d(i, s) = |\sum_{y < \ell(i, s)} \Phi_i^A(y)[s] \cdot 2^{-y} - \alpha_i[s]|,$$
$$md(i, s) = \min\{0, d(i, t) : t < s\}.$$

Say that stage $s$ is $i$-expansionary if $s = 0$ or if $\ell(i, s) > \max\{m(i, s), |\sigma| : \sigma$ is assigned to some $S_{i,j}$-strategy\}, $d(i, s) < \min\{md(i, s), 2^{-m\ell(i, s)}, 2^{-|\sigma|} : \sigma$ is assigned to some $S_{i,j}$-strategy\}, and $\alpha_i[s] \geq 2^{-F_i,t}$ for all $t < s$.

The $\tau$-module has a (finitary) outcome 1, in case that there are only finitely many $i$-expansionary stages, and an (infinitary) outcome 0, in case that there are infinitely many $i$-expansionary stages.

If $\tau$-module has an outcome 0, then $\tau$ is responsible to construct an auxiliary c.e. set of strings $U_\tau$ such that $U_\tau$ is noncomputable, and that $U_\tau$ presents $\alpha_\tau$. $U_\tau$ is constructed by $\tau$'s substrategies. Links from $\tau$ to these substrategies will be created and removed alternatively in the whole construction. We will also ensure that $\sum_{\sigma \in U_{\tau,s}} 2^{-|\sigma|} \leq \alpha_{\tau,s}$, and infinitely often, we can update $U_\tau$ such that $\sum_{\sigma \in U_{\tau,s}} 2^{-|\sigma|} = \alpha_{\tau,s}$. We can do this by Chaitin-Kraft inequality. Thus, we can effectively expand the prefix-free set $U_{\tau,s}$ to a larger one with the prefix-freeness preserved. Since $\alpha_i \in [0, \frac{1}{2}]$, we can always choose the strings for the updating process as $1r_\tau$, and leave the strings of the form $00 \cdots 01$ for the $S_{i,j}$-strategies, $j \in \omega$. This arrangement avoids the interaction between the diagonalization and the updating process.
(3.3) The $\eta$-module for $S_{i,j}$

The $\eta$-module is the “gap-cogap” argument, which is the key part of the proof. $\eta$-module requires various parameters such as $\sigma(\eta), r(\eta)$, whose values at the end of stage $s$ will be denoted by $\sigma(\eta)[s], r(\eta)[s]$. $\eta$-module has two objectives. One is to find some string $\sigma$ witnessing $U_\eta \neq \overline{V_\eta}$, and if this fails, $\eta$ will ensure in a definite way that the real $\alpha_\eta$ (namely, $\alpha_{r(\eta)}$) is computable. In the first case, $\eta$ itself is satisfied. In the second case, $R_i$ is satisfied and no further subrequirements $S_{i,j'}, j' \neq j$, need to be examined.

The $\eta$-module consists of the following steps:

1. Choose a string $\sigma = 00 \cdots 01$, with $|\sigma|$ big.

2. Wait for a stage $s$ such that $\sigma \in V_j[s]$. At stage $s + 1$ open a gap by setting $r[s + 1] = 0$, and go to step (3). Stop updating $U_i$ until this gap is closed.

3. Wait for the least $i$-expansionary stage $t \geq s + 1$ such that $\ell(i, t) > |\sigma|$. At stage $t + 1$ close the gap by performing step (3a) or (3b) according to which case applies:

   (3a) **(Successful close)** Suppose that $\alpha_i[t] - \alpha_i[s] > 2^{-|\sigma|}$. Enumerate $\sigma$ into $U_i$, update $U_i$ and stop.

   (3b) **(Unsuccessful close)** Otherwise, define $r[t + 1] = t$ (to preserve $\Phi_i^A[t] \upharpoonright |\sigma|$). Cancel $\sigma$, update $U_i$ and go to step (1).

Note that if $\alpha_i = \Phi_i^A$ and $F_i$ is a presentation of $\alpha_i$, then every gap must be eventually closed, and so $U_i$ will be updated correctly. $\eta$-module has three possible outcomes:

$w$. In this outcome, the $\eta$-module opens at most finitely many gaps and never closes one successfully.

Let $s$ be the last stage at which $\eta$ closes its gap, and $\sigma$ be chosen at stage $t > s$. Then $\sigma$ never goes into $V_j$, because otherwise, $\eta$ will create another link, which is contrary to the choice of $s$. Thus, $\overline{V_j}(\sigma) = 1 - V_j(\sigma) = 1 \neq 0 = U_i(\sigma)$. $S_{i,j}$ is satisfied.

$s$. There is some stage $s + 1$ at which $\eta$ closes a gap successfully.

In this case, $\sigma = \sigma[s]$ is enumerated into $U_i$ at stage $s + 1$. Suppose that at stage $v_1 < s$, $\sigma$ was defined as $\sigma[s]$, and at stage $v_2 \in [v_1, s)$, $\eta$ opened a gap trying to enumerate $\sigma$ into $U_i$. By the module, $\alpha_i[s] - \alpha_i[v_2] > 2^{-|\sigma|}$, since the updating process does not perform during the gap, we are eligible to put $\sigma$ into $U_i$. Thus, $U_i(\sigma) = 1 \neq 0 = 1 - \overline{V_j}(\sigma), S_{i,j}$ is satisfied.

$g$. In this outcome, $\eta$ opens infinitely many gaps and closes each one unsuccessfully.

Obviously, the limit of the length of the candidates (strings) is $\infty$. Suppose that the infinitely many gaps are begun at stages $\{t_n\}_{n \in \omega}$ and closed at stages $\{v_n\}_{n \in \omega}$, where $t_0 < v_0 < t_1 < v_1 < \cdots$. The intervals of stages $\{s : t_n \leq t < v_n\}$ are called gaps, since $r = 0$ in these intervals and we are free to enumerate any number into $A$ to make $A$ noncomputable, while the intervals $\{s : v_n \leq s < t_{n+1}\}$ are called cogaps and we have $r = v_n$.

$\alpha_i$ is computable because:
Given $n$. Choose a stage $t$ large enough such that a gap is begun at stage $t+1$, and the corresponding string $\sigma$ has length $|\sigma| > n + 3$. Since the gap is closed unsuccessfully at stage $v + 1 > t + 1$, we have $\alpha_i[v] - \alpha_i[t] \leq 2^{-|\sigma|}$. If the next gap is begun at stage $t' + 1 > v + 1$, then the $A$-restraint $r$ ensures that $\Phi^A \upharpoonright |\sigma|\uparrow = \Phi^A \upharpoonright |\sigma|[t']$. Thus,

$$
\alpha_i[t'] - \alpha_i[v] \\
\leq |\alpha_i[t'] - \sum_{y < \ell(i,v)} \Phi^A_i(y)[v] \cdot 2^{-y}] + |\sum_{y < \ell(i,v)} \Phi^A_i(y)[v] \cdot 2^{-y} - \alpha_i[v]| \\
\leq |\alpha_i[t'] - \sum_{y < \ell(i,v') \Phi^A_i(y)[t'] \cdot 2^{-y}] + |\sum_{y < \ell(i,v')} \Phi^A_i(y)[t'] \cdot 2^{-y} \\
- \sum_{y < \ell(i,v')} \Phi^A_i(y)[v] \cdot 2^{-y}] + |\sum_{y < \ell(i,v')} \Phi^A_i(y)[v] \cdot 2^{-y} - \alpha_i[v]| \\
\leq 2^{-|\sigma|} + 2^{-|\sigma|} + 2^{-|\sigma|} \\
= 3 \cdot 2^{-|\sigma|}.
$$

Therefore, we have

$$
\alpha_i[t'] - \alpha_i[t] = (\alpha_i[t'] - \alpha_i[v]) + (\alpha_i[v] - \alpha_i[t]) \\
\leq 3 \cdot 2^{-|\sigma|} + 2^{-|\sigma|} = 2^{-(|\sigma| - 2)}.
$$

Now the argument continues with $\sigma$ replaced by a new one, $\sigma'$ say. This establishes the following:

$$
\alpha_i - \alpha_i[t] \leq 2^{-(|\sigma| - 2)} + 2^{-(|\sigma'| - 2)} + \ldots \leq 2^{-(|\sigma| - 3)}.
$$

Since $|\sigma| > n + 3$, we have $\alpha_i - \alpha_i[t] < 2^{-n}$. Thus, $\alpha_i$ is computable.

In this case, $\liminf_s r[s] = 0$, because $r$ is defined as 0 during gaps.

$g$ is an infinitary outcome, while $w, s$ are both finitary.

## 4 The Priority Tree

Let $A = \{0, 1, s, g, w\}$ with the order $<_A$ in the order (from left to right) as listed. $T$ is defined as follows as a subset of the finite sequences of $A$.

### (4.1) The priority tree

1. Define $T = \{ \xi \in A^{<\omega} : (\forall n)[\xi(3n) \in \{0, 1\}, \xi(3n + 1) \in \{s, w\} & \xi(3n + 2) \in \{s, g, w\}] \}$.

2. Define $[T] = \{ h \in A^\omega : (\forall n)[h \upharpoonright n \in T] \}$, the set of all infinite path through $T$.

Every node $\tau \in T$ with $|\tau| \equiv 0 \pmod{3}$, will be assigned to the $R$-requirement. The outcome 1 for $\tau$ denotes that there are at most finitely many $\tau$-expansionary stages, so $\tau$ believes that $R$ is satisfied. The outcome 0 denotes that there are infinitely many $\tau$-expansionary stages, and that further actions to satisfy $R$ must be taken by nodes $\eta \supset \tau$.

Every node $\rho \in T$ with $|\rho| \equiv 1 \pmod{3}$, will be assigned to the $P$-requirement.

Every node $\eta \in T$ with $|\eta| \equiv 2 \pmod{3}$, will be assigned to some subrequirement $S_{i,j}$. The $\eta$-strategy will be a variant of the basic module described in the section above, having outcomes from $\{s, g, w\}$.

### (4.2) Define (partial) functions $i, j : T \to \omega$ and lists $L_0, L_1, L_2$, by induction on $n = |\xi|$ as follows.
n = 0. Define as above, \(i(\lambda) = 0\), and set \(L_0(\lambda) = L_1(\lambda) = L_2(\lambda) = \omega\).

\(n > 0\). Let \(\xi^-\) be the immediate predecessor of \(\xi\).

Assume that \(i(\xi^-), L_0(\xi^-), L_1(\xi^-), L_2(\xi^-)\) are defined, and that \(j(\xi^-)\) is defined if \(\xi^-\) is an \(S\)-strategy. First define \(L_0(\xi), L_1(\xi)\) and \(L_2(\xi)\) as those in the next paragraph. If \(\xi\) is an \(R\)-strategy, define \(i(\xi) = \mu i \in L_0(\xi)\), and let \(j(\xi)\) be undefined. If \(\xi\) is a \(P\)-strategy, define \(i(\xi) = \mu i \in L_1(\xi)\), and also let \(j(\xi)\) be undefined. If \(\xi\) is an \(S\)-strategy, define \(i(\xi) = i_0, j(\xi) = j_0\) where \(\langle i_0, j_0 \rangle = \mu n \in L_2(\xi)\).

Case 1: \(|\xi^-| \equiv 0 \pmod{3}\). Adopt the subcase below.

\[\quad\text{subcase (i): } a = 0. \text{ Define } L_0(\xi) = L_0(\xi^-) - \{i(\xi^-)\}, L_1(\xi) = L_1(\xi^-), \text{ and } L_2(\xi) = L_2(\xi^-).\]

\[\quad\text{subcase (ii): } a = 1. \text{ Define } L_0(\xi) = (L_0(\xi^-) - \{i(\xi^-)\}) \cup \{i': i' > i(\xi^-)\}, L_1(\xi) = L_1(\xi^-) \cup \{i': i' > i(\xi^-)\}, \text{ and } L_2(\xi) = (L_2(\xi^-) - \{i(\xi^-), k : k \in \omega\}) \cup \{i', k : i' > i(\xi^-) \& k \in \omega\}.\]

Case 2: \(|\xi| \equiv 1 \pmod{3}\). Define \(L_0(\xi) = L_0(\xi^-), L_1(\xi) = L_1(\xi^-) - \{i(\xi^-)\}, \text{ and } L_2(\xi) = L_2(\xi^-).\)

Case 3: \(|\xi| \equiv 2 \pmod{3}\). Adopt the subcase below.

\[\quad\text{subcase (i): } a = s \text{ or } a = w. \text{ Define } L_0(\xi) = L_0(\xi^-), L_1(\xi) = L_1(\xi^-), \text{ and } L_2(\xi) = L_2(\xi^-) - \{i(\xi^-), j(\xi^-)\}.\]

\[\quad\text{subcase (ii): } a = g. \text{ Define } L_0(\xi), L_1(\xi) \text{ and } L_2(\xi) \text{ as in Case 1, subcase (ii).}\]

This completes the priority assignment.

## 5 Construction

At the end of the construction, we define the true path \(f \in [T]\) of the construction. \(f\) is approximated via \(\delta[s] \in T\) at each stage \(s\), and we will have \(f = \lim \inf_s \delta[s]\). During stage \(s\), we define \(\delta[s]\) via a series of substages \(t < s\) at which we define a string \(\delta_t\) such that \(\delta_t \subseteq \delta_{t+1}\) and \(\delta[s]\) is the last \(\delta_t\) defined. Only string \(\delta_t\) is allowed to act at substage \(t\).

### (5.1) Genuine stages
For \(\xi \in T\),

1. Say that \(s + 1\) is a \(\xi\)-stage if \(\xi \subseteq \delta[s + 1]\). In addition 0 is a \(\xi\)-stage. Let \(S^\xi\) be the set of all \(\xi\)-stages.

2. Say that \(s + 1\) is a genuine \(\xi\)-stage if \(\xi = \delta_t\) for some substage \(t\) of stage \(s + 1\), where \(\delta_t\) will be defined during the construction. Let \(G^\xi\) be the set of genuine \(\xi\)-stages.

### (5.2) Expansionary stages
Fix \(\tau\). At stage \(s + 1\), define the length functions as:

\[\ell(\tau, s) = \max\{x < s : (\forall y < x)[\Phi^A_\tau(y)[s] \downarrow}\},\]

\[m\ell(\tau, s) = \max\{0, \ell(\tau, t) : t \leq s \& t + 1 \in S^\tau\}.\]

Also define the distance function as:

\[d(\tau, s) = |\sum_{y < \ell(\tau, s)} \Phi^A_\tau(y)[s] \cdot 2^{-y} - \alpha_\tau[s]|,\]

\[md(\tau, s) = \min\{0, d(\tau, t) : t \leq s \& t + 1 \in S^\tau\}.\]
Say that stage \( v \) is \( \tau \)-expansionary stage, if \( v = 0 \), or if \( v = s + 1 \in S^\tau \), \( \ell(\tau, s) > \max\{ml(\tau, s), |\sigma| : \sigma \text{ is ever assigned to some } S_{i(\tau), j}-\text{strategy}\} \), \( d(\tau, s) < \min\{md(\tau, s), 2^{-m\ell(\tau, s)}, 2^{-|\sigma|} : \sigma \text{ is ever assigned to some } S_{i(\tau), j}-\text{strategy}\} \), and \( \alpha_\tau[s] > 2^{-F_{\tau,t}} \) for all \( t < s, t \in S^\tau \).

Suppose that at stage \( s \), an \( S \)-strategy \( \eta \) opens a gap, then we shall construct a link \((\tau, \eta)\) with bottom \( \eta \) and top \( \tau = \tau(\eta) \). By the construction of \( T \), we have \( \tau \sim 0 \) \( \subseteq \eta \) and \( F(\eta)[s] = g \). The link \((\tau, \eta)\) remains in force until either \( \eta \) closes the gap (either a successful or a unsuccessful close), or \( \eta \) is initialized. If at stage \( s + 1 \), there is a link \((\tau, \eta)\) and there is a substage \( t \) of stage \( s + 1 \) such that \( \delta_t = \tau \) and \( \delta_{t+1} = \eta \), then we say that the link \((\tau, \eta)\) is travelled at stage \( s + 1 \). The purpose of the link is to allow the flow of control of the construction to pass directly from the top node \( \tau \) to the bottom node \( \eta \) at \( \tau \)-expansionary stages, to close the gap. The hypothesis \( \alpha_\tau = \Phi_\tau(A) \) guarantees that each link will be travelled (at most once) and the corresponding gap will be closed.

Note that travelling the link \((\tau, \eta)\) may lead the construction to pass over some nodes \( \xi, \tau \subset \xi \subset \eta \) of lower global priority, i.e., \( i(\xi) > i(\eta) = i(\tau) \), which would otherwise receive attention before \( \eta \) if we simply descended the tree without using links. For such \( \xi \), \( s + 1 \) is a \( \xi \)-stage, but not a genuine \( \xi \)-stage. This can result in the failure of \( \xi \). Thus, to consider whether a given \( \xi \)-strategy succeeds or not, we must examine the genuine \( \xi \)-stages rather than merely the \( \xi \)-stages.

(5.3) (Construction)

The construction is as follows:

Stage \( s = 0 \). Initialize all nodes \( \xi \in T \). Define \( \delta[0] = \lambda \).

Stage \( s + 1 \). The construction will proceed by substages \( t \leq s + 1 \). We refer to substage \( t \) of stage \( s + 1 \) as stage \((s + 1, t)\). The value of parameter \( p \) at the end of substage \( t \) will be denoted by \( p_t \), rather that the complete specification \( p_t[s + 1] \). At the end of substage \( t \) we shall have \( \delta_t \in T \) such that \( \delta_t \subseteq \delta_{t+1} \subseteq \delta[s + 1] \), for \( t \leq s \), and only \( \delta_t \) may act during substage \( t \). At the end of stage \( s + 1 \), if \( s + 1 \) is a \( \tau \)-expansionary stage, and there is no link with top \( \tau \) created at stage \( s + 1 \), where \( \tau \) is an \( R \)-strategy, then perform the updating process for \( \tau \) by enumerating into \( U_\tau \) appropriate strings to make \( \sum_{\sigma \in U_\tau} 2^{-|\sigma|} = \alpha_{\tau,s+1} \). This process keeps the resulted c.e. set \( U_{\tau,s+1} \) prefix-free.

Substage \( t = 0 \). Define \( \delta_0 = \lambda \). Go to substage 1.

Substage \( t + 1 \leq s + 1 \). Given \( \delta_t \) and for all \( S \)-strategies \( \eta \in T \), \( F_t(\eta) \in \{s, g, w\} \), the current state of the \( \eta \)-module, define \( \delta_{t+1} \) as follows:

Case 1. \( \delta_t \) is an \( R \)-strategy. If \( s + 1 \) is not \( \delta_t \)-expansionary, then define \( \delta_{t+1} = \delta_t \sim 1 \). If \( s + 1 \) is \( \delta_t \)-expansionary, and there is no link with top \( \delta_t \), then define \( \delta_{t+1} = \delta_t \sim 0 \). Otherwise, \( s + 1 \) is \( \delta_t \)-expansionary, and there is a link with top \( \delta_t \) and bottom \( \eta \supseteq \delta_t \sim 0 \), define \( \delta_{t+1} = \eta \).

Case 2. \( \delta_t \) is a \( P \)-strategy. If \( \delta_t \) is satisfied, define \( \delta_{t+1} = \delta_t \sim 0 \). Otherwise, define \( \delta_{t+1} = \delta_t \sim 1 \).

Case 3. \( \delta_t \) is an \( S \)-strategy. Define \( \delta_{t+1} = \delta_t \sim F_t(\delta_t) \).

Let \( \xi = \delta_{t+1} \). Say that \( \xi \) requires attention at substage \((s + 1, t + 1)\) if \( \xi \) has not been satisfied, and:}
1. $\xi$ is a $P$-strategy, and one of the following holds:
   (1a) $x(\xi)$ is not defined, or
   (1b) $\Phi_{\xi}(x(\xi))[s] \downarrow$.

2. $\xi$ is an $S$-strategy, and one of the following holds:
   (2a) $\sigma(\xi)$ is not defined, or
   (2b) $\xi$ is ready to open a gap, or
   (2c) $\xi$ is ready to close a gap.

If $\xi$ does not require attention, and $t < s$, then go to substage $t+2$. Otherwise, choose the first one which applies and perform the action as follows. If $t < s$, and $\xi$ opens an $\xi$-gap, then go to substage $t + 2$. Otherwise, define $\delta[s+1] = \delta_{t+1}$, go to stage $s+2$.

**Case 1**: $\xi = \rho$ is a $P$-strategy. If (1a) happens, then define $x(\rho)$ as a big number. If (1b) happens, then enumerate $x(\rho)$ into $A$ and say that $P_\rho$ is satisfied. In any case, initialize all $\zeta > \rho$.

**Case 2**: $\xi = \eta$ is an $S$-strategy.

If (2a) happens, define $\sigma(\eta) = 00 \cdots 01$ with $|\sigma(\eta)|$ big. (Clearly, $\{\sigma(\eta)\} \cup U_{\tau(\eta),s}$ is prefix-free.) Initialize all $\zeta \geq \eta \wedge w$.

If (2b) happens (This action is called step 1), open an $\eta$-gap by defining $F(\eta) = g$, and $r(\eta) = 0$. Initialize all $\zeta \geq \eta \wedge w$. Create a link $(\tau,\eta)$ with top $\tau = \tau(\eta)$, and bottom $\eta$.

If (2c) happens (This action is called step 2), let $v + 1 < s + 1$ be the stage at which the current $\eta$-gap was opened, and let $\sigma = \sigma(\eta)[s]$. Close the $\eta$-gap as follows:

**step 2a (Successful Close)** If $\alpha_\eta[s] - \alpha_\eta[v] \geq 2^{-|\sigma(\eta)|}$, then enumerate $\sigma(\eta)$ into $U_{\tau(\eta)}$, and define $F(\eta) = s$. Initialize all $\zeta \geq \eta \wedge w$. Set $r(\eta) = 0$.

**step 2b (Unsuccessful Close)** If $\alpha_\eta[s] - \alpha_\eta[v] < 2^{-|\sigma(\eta)|}$, then define $F(\eta) = w$ and $r(\eta) = s$ (to preserve $\Phi_\eta \upharpoonright |\sigma(\eta)|$), cancel $\sigma(\eta)$, initialize all $\zeta \geq \eta \wedge w$.

This completes the construction.

### 6 Verification

Let $f$ be the leftmost path. That is, the leftmost $f \in [T]$ such that for all strings $\xi \in T$, if $\xi \leq_L f$ and $\xi \not\subseteq f$, then there are only finitely many stages where $\xi \subseteq \delta[s]$. The following lemmas prove that our construction satisfies all requirements.

**Lemma** Fix $n$ and $\xi \subset f$ with $|\xi| = n$. Then,

1. $(\exists s)(\xi \not\subseteq \delta_s)$,
2. $|G^\xi| = \infty$,
3. If $|\xi| \equiv 1 \pmod{3}$, then $\xi$ receives attention only finitely often.
4. For any $\tau \subset \xi$, if $|\tau| \equiv 0 \pmod{3}$, and $\tau \wedge \langle 0 \rangle \subset \xi$, then every link with top $\tau$ is removed and there are infinitely many genuine $\xi$-stages where there are no links with top $\tau$. 

(6.2) Lemma Let $\rho \subset \xi$ with $|\rho| \equiv 1 \pmod{3}$, and let $x$ be a follower of $\rho$ where $x = \lim_{s} x(\rho)[s]$, the last follower of $\rho$ appointed at stage $s_1$ say, where $W_{e,s_1} \cap A_{s_1} = \emptyset$. Then

1. $(\forall s \in G^\rho) [s > s_1 \rightarrow y > R(\rho,s)]$, where $R(\rho,s) = \max \{r(\zeta,s) : \zeta \leq L \rho \}$.
2. $P_i(\rho)$ is satisfied.

(6.3) Lemma Suppose that $\tau \subset f$ with $|\tau| \equiv 0 \pmod{3}$ and $\tau \prec \langle 1 \rangle \subset f$. Then

$\alpha_{i(\tau)} \neq \Phi_{i(\tau)}^{A}$.

(6.4) Lemma Suppose that $\eta \subset f$ with $|\eta| \equiv 2 \pmod{3}$.

1. If $\eta \prec \langle s \rangle \subset f$, then $U_{\tau(\eta)} \cap V_{j(\eta)} \neq \emptyset$. Indeed, $\sigma(\eta) \in U_{\tau(\eta)} \cap V_{j(\eta)}$, where $\sigma(\eta)$ is the (unique) string enumerated into $U_{\tau(\eta)}$ when the (last) $\eta$-gap is closed (successfully).
2. If $\eta \prec \langle g \rangle \subset f$, then $\alpha_{\eta}$ is computable.
3. If $\eta \prec \langle w \rangle \subset f$, then $U_{\tau(\eta)} \cup V_{j(\eta)} \neq \Sigma^*$.

(6.5) Lemma Suppose that $\tau \subset f$ with $|\tau| \equiv 0 \pmod{3}$ and $\tau \prec \langle 0 \rangle \subset f$. Then $U_{\tau}$ is a presentation of $\alpha_{\tau}$. Namely, $U_{\tau}$ is prefix-free, and $\sum_{\sigma \in U_{\tau}} 2^{-|\sigma|} = \alpha_{\tau}$. Moreover, if $\alpha_{\tau}$ is not computable, then $U_{\tau}$ is also not computable.

This completes the proof of Theorem 6.

References

Weakly Computable Real Numbers and Total Computable Real Functions (Extended Abstract)

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Abstract. Let $C_{sc}$ and $C_{wc}$ be classes of the semi-computable and weakly computable real numbers, respectively, which are discussed by Weihrauch and Zheng [12]. In this paper we show that both $C_{sc}$ and $C_{wc}$ are not closed under the total computable real functions of finite length on some closed interval, although such functions map always a semi-computable real numbers to a weakly computable one. On the other hand, their closures under general total computable real functions are the same and are in fact an algebraic field. This field can also be characterized by the limits of computable sequences of rational numbers with some special converging properties.

1 Introduction

The computability notion of real numbers is introduced by Alan M. Turing in his famous paper [9] on “Turing machine”. He called a real number $x$ computable if it has a computable binary expansion. Namely, there is a recursive set $A \subseteq \mathbb{N}$ such that $x = x_A := \sum_{i \in A} 2^{-(i+1)}$. Equivalently, $x$ is computable iff there is a computable sequence $(x_n)_{n \in \mathbb{N}}$ of rational numbers which converges to $x$ effectively in the sense that $|x - x_n| \leq 2^{-n}$ for all $n \in \mathbb{N}$. In fact, as pointed by R. M. Robinson [7], corresponding to any classical definition of real numbers there is an equivalent definition of computable real numbers. The class of all computable real numbers is denoted by $C_e$.

Essentially, we can effectively find an upper bound and a lower bound of any computable $x \in \mathbb{R}$ within any given error. In other words, we can approximate the real number $x$ effectively from both sides, below and above. Sometimes, a real number can only be approximated effectively from one side instead of two. As one of the first such example, Specker [8] has shown that the real number $x_A$ for a nonrecursive r.e. set $A$ is effectively approximable from below but it is not effectively approximable from above. Such kind of real numbers are called left computable$^1$. Namely, a real number is left computable if it is a limit of

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$^1$ Some authors call them recursively enumerable or computably enumerable (see, e.g. [2]).
some computable increasing sequence of rational numbers. The \textit{right computable real numbers} can be defined accordingly. Left and right computable real number classes are denoted by $C_{lc}$ and $C_{rc}$, respectively.

It is not difficult to see that $C_e$ is closed under the arithmetic operations, namely, it forms an algebraic field. However, the situation for $C_{lc}$ and $C_{rc}$ is different. The first easy observation is, that $x$ is left computable iff $-x$ is right computable, and $x$ is computable iff it is both left and right computable. From Specker’s example, it follows immediately that the classes $C_{lc}$ and $C_{rc}$ are not closed under the arithmetic operations, hence they are not algebraic fields. Moreover, Weihrauch and Zheng [12] have shown that there is a left computable real number $x$ and a right computable real number $y$ such that the sum $x+y$ is neither left computable nor right computable. Therefore even the union $C_{sc} := C_{lc} \cup C_{rc}$ (the class of \textit{semi-computable} real numbers) is not an algebraic field. In other words, the closure of semi-computable real numbers under the arithmetic operations is a proper superset of the semi-computable real number class. This closure is denoted by $C_{wc}$ and any real number in this closure is called \textit{weakly computable} (w.c., for short). An interesting characterization of w.c. real numbers is (see [12]), that $x$ is weakly computable iff there is a computable sequence $(x_n)_{n \in \mathbb{N}}$ of rational numbers which converges to $x$ \textit{weakly effectively} in the sense that $\sum_{n \in \mathbb{N}} |x_n - x_{n+1}|$ is finite.

As the weakest version of effectiveness on real numbers we can consider simply the limits of any computable sequence of rational numbers without any further restriction but the convergence. We will call a real number $x$ \textit{recursively approximable} (r.a., for short) if there is a computable sequence of rational numbers $(x_n)_{n \in \mathbb{N}}$ of rational numbers which converges to $x$. The class of all r.a. real numbers is denoted by $C_{ra}$. As shown by Weihrauch and Zheng [12], the class $C_{wc}$ is a proper subset of $C_{ra}$ and $C_{wc}$ is an algebraic field\(^2\).

The class $C_{ra}$ shares a lot of mathematical properties with $C_e$. For example, they are all closed under the computable operations (i.e. partial computable real functions) and effectively convergent limits (13). On the other hand, the class $C_{wc}$ does not share these properties any more, as shown by Zheng [13], although it is also an algebraic field. One of the interesting problems remaining open there is, whether $C_{sc}$ or $C_{wc}$ is closed under the total computable real functions. A negative answer to this question will be shown in this paper.

During the discussion of closure properties for real number classes under the total computable real functions, it plays an important role, that whether a total computable real function has a finite length on some closed interval\(^3\). For example, we can easily show that the image $f(x)$ of a semi-computable real number $x$ under a computable real function $f$ of finite length on some closed interval containing $x$ is always weakly computable. But $f(x)$ can be non-weakly computable for general total computable real functions $f$. Furthermore, we show that the class $C_{sc}$ is not closed under the total computable real functions, no matter the length of such functions. This holds also for the class $C_{wc}$. In fact

\[^2\] An earlier but weaker result that $C_{sc} \subset C_{ra}$ is proved by Ceit in [3].

\[^3\] The length of the a function means here the length of its graph.
we can even show that $C_{wc}$ is not closed under the monotone computable total real functions. On the other hand the closures of $C_{wc}$ and $C_{lc}$ under general total computable real functions are in fact the same. This common closure is an algebraic field and is strictly between the classes $C_{wc}$ and $C_{ra}$. At the end we show that this field can also be characterized by computable sequences of rational numbers with some restriction on their convergence.

2 Preliminaries

In this section we will recall some notions and related known results. Basic notions about classical computability theory are assumed. Since the real numbers can be described by rational numbers, the effectiveness notion on real numbers will be naturally introduced by means of the effectiveness notions on rational numbers. On rational numbers $\mathbb{Q}$, the computability theory is well developed by classical computability theory. For example, the notions of computable function $f : \mathbb{Q} \to \mathbb{Q}$, recursively enumerable set $A \subseteq \mathbb{Q}$, etc. are clear and do not need any further explanation. Especially, we will call a sequence $(x_n)_{n \in \mathbb{N}}$ of rational numbers computable, if there is a computable function $f : \mathbb{N} \to \mathbb{Q}$ such that $x_n = f(n)$.

In the last section we have mentioned several classes of real numbers. The relationship among these classes looks like, as shown in [12], the following

$$C_e = C_{lc} \cap C_{rc} \subsetneq C_{lc} \subsetneq C_{rc} \subsetneq C_{sc} = C_{lc} \cup C_{rc} \subsetneq C_{wc} \subsetneq C_{ra}.$$ 

For $C_{wc}$, we have the following equivalent descriptions.

**Theorem 1** (Weihrauch and Zheng [12]). For $x \in \mathbb{R}$, the following are equivalent:

1. $x$ is weakly computable;
2. There are left computable real numbers $y, z \in \mathbb{R}$ such that $x = y - z$;
3. $x$ is in the arithmetic closure of $C_{lc}$;
4. There is a computable sequence $(x_n)_{n \in \mathbb{N}}$ of real numbers which converges to $x$ such that $\sum_{n=0}^{\infty} |x_n - x_{n+1}| \leq c$ for some real number $c$.

The computability notion for real numbers can be extended to sequences. A sequence $(x_n)_{n \in \mathbb{N}}$ of real numbers is called *computable*, if there is a computable sequence $(r_{nm})_{n,m \in \mathbb{N}}$ of rational numbers which converges to $(x_n)_{n \in \mathbb{N}}$ uniformly effectively in $n$ and $m$. Namely, for any $n, m \in \mathbb{N}$, $|r_{nm} - x_n| \leq 2^{-m}$ holds.

Now let us recall the definition of computable real functions. There are a lot of equivalent approaches to define the computable real functions (see e.g. [4][5][6][11]). Here we use the approach of Weihrauch [10][11] by so-called Type-2 Turing machines. A type-2 Turing machine $M$ extends the classical Turing machine in such a way that it accepts infinite sequences as well as finite strings as inputs and outputs. For any input $p$, $M(p)$ outputs a finite string $q$ if $M(p)$ halts in finite steps with $q$ in its write-only output tape, while $M(p)$ outputs
an infinite sequence \( q \) means that \( M(p) \) will never halt and keep writing the sequence \( q \) on its output tape. A real function\(^1\) \( f : \subseteq \mathbb{R} \to \mathbb{R} \) is \textit{computable}, if there is a type-2 Turing machine \( M \) which computes \( f \) in the sense that, for any \( x \in \text{dom}(f) \) and any sequence \( p \) (of rational numbers) which converges effectively to \( x \), \( M(p) \) outputs a sequence \( q \) (of rational numbers) which converges effectively to \( f(x) \). It is easy to see that computable function is continuous at any point of its domain.

In this paper we are mainly interested in the closure properties of some real number classes under the computable total real functions. Obviously, the real number \( x \) and \( x \pm n \), for any \( n \in \mathbb{N} \), have completely the same effectiveness in any reasonable sense. Therefore, w.l.o.g., we consider only the real numbers in the interval \([0; 1]\) which can always be expressed by \( x_A \) for some \( A \subseteq \mathbb{N} \). Similarly, we can restrict ourselves to the computable real functions \( f : [0; 1] \to [0; 1] \) for which we have the following equivalent descriptions.

\textbf{Theorem 2 (cf. Pour-El and Richards \cite{PeRi89}, Weihrauch \cite{Wei96}).}
\begin{itemize}
  \item 1. \( f \) is computable;
  \item 2. There is a computable sequence \((p_n)_{n \in \mathbb{N}}\) of rational polygon functions which converges to \( f \) uniformly effectively. Namely, \( |f(x) - p_n(x)| \leq 2^{-n} \) for any \( x \in [0; 1] \) and \( n \in \mathbb{N} \).
  \item 3. \( f \) is sequentially computable, i.e. \( f \) maps any computable sequence \((x_n)_{n \in \mathbb{N}}\) of real number to a computable sequence \((f(x_n))_{n \in \mathbb{N}}\); and \( f \) is effectively uniformly continuous, i.e. there is a recursive modulus \( e : \mathbb{N} \to \mathbb{N} \) such that \( \forall x, y \in [0; 1] \forall n \in \mathbb{N} \) \(|x - y| \leq 2^{-e(n)} \implies |f(x) - f(y)| \leq 2^{-n}\).
\end{itemize}

Item 2. of Theorem 2 is called Effective Weierstrass Theorem which implies immediately the following lemma where \([0; 1]_\mathbb{Q} := \{ r \in \mathbb{Q} : 0 \leq r \leq 1 \}\).

\textbf{Lemma 1.} Let \( f : [0; 1] \to \mathbb{R} \) be computable. There is a computable function \( \beta : \mathbb{N} \times [0; 1]_\mathbb{Q} \to \mathbb{Q} \) such that \( |f(r) - \beta(n, r)| \leq 2^{-n} \) for any \( n \in \mathbb{N} \) and \( r \in [0; 1]_\mathbb{Q} \).

In this case we call \( \beta \) an effective determinator of function \( f \).

Let \( \mathbf{C} \subseteq \mathbb{R} \) and \( \mathbf{F} \) a set of functions \( f : \subseteq [0; 1] \to \mathbb{R} \). Denote by \( \mathbf{F}(\mathbf{C}) \) the image set of \( \mathbf{C} \) under the functions of \( \mathbf{F} \). Furthermore we define \( \mathbf{F}^n(\mathbf{C}) \), for any \( n \in \mathbb{N} \), inductively by \( \mathbf{F}^0(\mathbf{C}) := \mathbf{C} \) and \( \mathbf{F}^{n+1}(\mathbf{C}) := \mathbf{F}(\mathbf{F}^n(\mathbf{C})) \). The closure \( \mathbf{F}^* \) of \( \mathbf{C} \) under the functions of \( \mathbf{F} \) is defined now by \( \mathbf{F}^*(\mathbf{C}) := \bigcup_{n \in \mathbb{N}} \mathbf{F}^n(\mathbf{C}) \). We are especially interested in the classes \( \text{FIN} \) and \( \text{TOT} \) of computable total real functions from \([0; 1]\) to \([0; 1]\) of finite and arbitrary length, respectively. For convenience we will call the function \( f \in \text{TOT} \) (\( f \in \text{FIN} \)) total computable real function (of finite length) without mention of the interval \([0; 1]\). Since the composition of two computable real functions is a computable real function, we have \( \text{TOT}^* = \text{TOT}^n = \text{TOT}^* \) for any \( n \geq 1 \). However \( \text{FIN}(\mathbf{C}) = \text{FIN}^*(\mathbf{C}) \) does not hold in general.

\(^1\) We use \( f : \subseteq X \to Y \) to denote that \( f \) is a partial function from \( X \) to \( Y \) while \( f : X \to Y \) is a total function from \( X \) to \( Y \).
Let \((M_e)_{e \in \mathbb{N}}\) be an effective enumeration of all Turing machines and \((\varphi_e)_{e \in \mathbb{N}}\) an effective enumeration of all computable functions \(\varphi_e : \mathbb{N} \to \mathbb{N}\) such that \(\varphi_e\) is computed by \(M_e\). We write \(\varphi_e(n) \downarrow\) if \(n \in \text{dom}(\varphi_e)\) and \(\varphi_e(n) \uparrow\), otherwise. The uniformly effective approximation \(\varphi_{e,s}\) of \(\varphi_e\) is defined by \(\varphi_{e,s}(n) := m\), if \(M_e(n)\) halts in \(s\) steps and outputs \(m\); and \(\varphi_{e,s}(n)\) is undefined, otherwise. Naturally, we assume that \(e, n, m \leq s\) if \(\varphi_{e,s}(n) = m\). For other types of the functions, say, \(f : \mathbb{N} \to \mathbb{Q}\), a similar effective enumeration and corresponding approximation can be introduced accordingly.

For any alphabet \(\Sigma\), let \(\Sigma^*\) and \(\Sigma^\omega\) be the sets of all finite strings and infinite sequences of \(\Sigma\), respectively. For \(u, v \in \Sigma^*\), denote by \(uv\) the concatenation of \(v\) after \(u\). If \(w \in \Sigma^* \cup \Sigma^\omega\), then \(w[n]\) denotes its \(n\)-th element. Thus, \(w = w[0]w[1] \cdots w[n - 1]\), if \(|w|\), the length of \(w\), is \(n\), and \(w = w[0]w[1]w[2] \cdots\), if \(|w| = \infty\). Obviously, \(w[n]\) is defined only for \(n < |w|\). We will say also that \(w[n]\) is undefined (written \(w[n] = \uparrow\)), if \(n \geq |w|\). The unique string of length 0 is always denoted by \(\lambda\) (so-called empty string). For any finite string \(w \in \{0, 1\}^*\), and number \(n \leq |w|\), the restriction \(w \upharpoonright n\) is defined by \((w \upharpoonright n)[i] := w[i]\) if \(i < n\) and \((w \upharpoonright n)[i] := \uparrow\), otherwise. The generalized restriction \(w \upharpoonright n\) is defined by \(w[n] := w \upharpoonright (n + 1)\), if \(n < |w|\) and \(w[n] := w(1^{(n+1)} - |w|)\), otherwise. For \(u, v \in \Sigma^* \cup \Sigma^\omega\), if \(u = v \upharpoonright n\) for some \(n \leq |v|\), then we call \(u\) an initial segment of \(v\) and denote it by \(u \sqsubseteq v\). We denote also \(u \neq v \& u \sqsubseteq v\) by \(u \sqsubseteq v\). Let \(\langle \cdot, \cdot \rangle : \mathbb{N}^2 \to \mathbb{N}\) be a computable pairing function with computable reverse functions.

### 3 Technical Lemma

In this section we will show two technical results which will be used to construct some total computable real functions satisfying some special conditions. We introduce at first a notion of interval trees.

**Definition 1.** Let \(\delta : \mathbb{N} \to \mathbb{N}^+\) be a function and \(\mathbb{I}\) the set of all rational intervals on \([0; 1]\) and \(\mathbb{N}_0^\delta := \{w \in \mathbb{N}^* : \forall n < |w|\ (w[n] < \delta(n))\}\).

1. A \(\delta\)-interval tree (\(\delta\)-i.t., for short) on \([0; 1]\) is a function \(I : \mathbb{N}_0^\delta \to \mathbb{I}\) such that \(I(\lambda) = [0; 1] \cup \bigcup_{i < \delta(|w|)} I(w[i]) = I(w)\), for any \(w \in \mathbb{N}_0^\delta\); and \(\lim_{s \to \infty} l(I(w_s)) = 0\) for any increasing sequence \((w_s)_{s \in \mathbb{N}}\) of \(\mathbb{N}_0^\delta\) in the sense of \(w_s \sqsubseteq w_{s+1}\), where \(l(I(w))\) is the length of the interval \(I(w)\). For any \(\delta\)-i.t. \(I\) and \(w \in \mathbb{N}_0^\delta\), the interval \(I(w)\) is denoted by \(I(w) := [a_w^\delta; b_w^\delta]\) and its length is denoted by \(l_w^\delta := b_w^\delta - a_w^\delta\).

2. A \(\delta\)-i.t. \(I\) is called computable if the functions \(a, b : \mathbb{N}_0^\delta \to [0; 1]_Q\) defined by \(a(w) := a_w^\delta\) and \(b(w) := b_w^\delta\), respectively, are computable.

3. A \(\delta\)-i.t. \(I\) is called canonical if, for any \(w \in \mathbb{N}_0^\delta\), the interval \(I(w)\) is divided into \(\delta(|w|)\) disjunctive subintervals \(I(w0), I(w1), \ldots, I(w(\delta(|w|) - 1))\) of the same length, namely, \(a_w^\delta := \sum_{i < |w|} w[i] \cdot \prod_{j \leq i} \delta(j)^{-1}\) and \(b_w^\delta := a_w^\delta + \prod_{j < |w|} \delta(j)^{-1}\) where \(\sum_{i \in \emptyset} (\cdots) := 0\) and \(\prod_{i \in \emptyset} (\cdots) := 1\).
Theorem 3. Notice that, for a canonical $\delta$-i.t. $I$, the set $S^I_w := \{ I(wi) : i < \delta(|w|) \}$ is a partition of the interval $I(w)$ and all subintervals in $S^I_w$ have the same length. But, general $\delta$-i.t. $I$, $S^I_w$ is not necessarily disjunctive. Besides, it is easy to see that a canonical $\delta$-i.t. is always computable, if $\delta$ is computable. We are interested mainly in $\delta$-i.t.'s on $[0; 1]$. So the phrase “on $[0; 1]$” is often omitted later. $\delta$-i.t.'s will be very useful to define some total computable real functions which satisfy special properties, if these properties can be approximated by some “compatible” function $i$ defined as follows.

Definition 2. Let $\delta_1, \delta_2 : \mathbb{N} \to \mathbb{N}^+$ and $i : \subseteq \mathbb{N}_{\delta_1}^* \to \mathbb{N}_{\delta_2}^*$ be functions.

1. $i$ is called $(\delta_1, \delta_2)$-compatible if the following conditions hold.
   (a) The domain $\text{dom}(i)$ of $i$ is infinite and is alternate in the sense that, $w(i - 1), w(i + 1) \not\in \text{dom}(i)$ for any $wi \in \text{dom}(i)$ and $0 < i < \delta_1(|w|)$;
   (b) $\forall w, v \in \mathbb{N}_{\delta_1}^* (w \in \text{dom}(i) \& v \subseteq w \implies v \in \text{dom}(i) \& i(v) \subseteq i(w))$;
   (c) $\forall u, v \in \text{dom}(i) (|u| = |v| \implies |i(u)| = |i(v)|)$.
2. $i$ is called $(\delta_1, \delta_2)$-increasing, if $b_{i(wj)}^{\delta_2} \leq a_{i(wj)}^{\delta_2}$ for any $w \in \text{dom}(i)$ and $i < j$.
3. $i$ is called effectively modular if there is a recursive function $g : \subseteq \mathbb{N} \to \mathbb{N}$ such that, $wi \in \text{dom}(i)$ iff $i \in \text{dom}(g) \& i(wi) = i(w)g(i)$, for any $w \in \text{dom}(i)$.

Theorem 3. Let $\delta_1, \delta_2 : \mathbb{N} \to \mathbb{N}^+$ be recursive functions, $I_1$ a canonical $\delta_1$-i.t., and $I_2$ a computable $\delta_2$-i.t. such that $b_{\delta_2}^{\delta_1} \leq 2^{-c(|w|)}$ for some recursive function $c$ and all $w \in \mathbb{N}_{\delta_2}^*$. If $i : \subseteq \mathbb{N}_{\delta_1}^* \to \mathbb{N}_{\delta_2}^*$ is a $(\delta_1, \delta_2)$-compatible computable function, then there is a computable function $f : [0; 1] \to [0; 1]$ such that $f(I_1(w)) \subseteq I_2(i(w))$ for all $w \in \text{dom}(i)$. If $i$ is also $(\delta_1, \delta_2)$-increasing, then such computable function $f$ can even be also increasing.

Notice that, a function $f$ is increasing here means that $x \leq y \implies f(x) \leq f(y)$, namely non-strictly increase. The computable function $f$ constructed in Theorem 3 can have finite length if it is increasing. Next Theorem shows that $f$ can have also finite length even if it is not monotone.

Theorem 4. Let $\delta_1 := k_1$ and $\delta_2 := k_2$ be constant functions, $I_1$ and $I_2$ the canonical $\delta_1$-i.t. and $\delta_2$-i.t., respectively. If $2k_2 \geq k_1$ and $i : \subseteq \mathbb{N}_{\delta_1}^* \to \mathbb{N}_{\delta_2}^*$ is a $(\delta_1, \delta_2)$-compatible and effectively modular function, then there is a computable function $f : [0; 1] \to [0; 1]$ of finite length such that $f(I_1(w)) \subseteq I_2(i(w))$ for all $w \in \text{dom}(i)$.

4 Closure Properties under FIN

In this section we discuss the closure properties of some real number classes under total computable real functions of finite length. We will see that neither $\mathbb{C}_{sc}$ nor $\mathbb{C}_{wc}$ is closed under FIN.

Notice that, the computable function $f : [0; 1] \to [0; 1]$ defined by $f(x) := 1-x$ has finite length and maps left computable real numbers to right computable ones and vice versa. Furthermore, for any $g \in \text{FIN}$, $g \circ f$ and $g$ have the same
length. Then it is easy to see that \( \text{FIN}(C_{sc}) = \text{FIN}(C_{rc}) = \text{FIN}(C_{wc}) \). Therefore we often state our results only about the class \( \text{FIN}(C_{sc}) \) and omit the other two classes. We show at first that \( C_{sc} \) is not closed under \( \text{FIN} \) and its image \( \text{FIN}(C_{sc}) \) is contained in \( C_{wc} \).

**Theorem 5.** \( C_{sc} \subseteq \text{FIN}(C_{sc}) \subseteq C_{wc} \).

**Proof.** (Sketch for the first part) We define a computable function \( f : [0; 1] \to [0; 1] \) by applying Theorem 4. Let \( \delta(n) = \delta_2(n) = 7 \) for any \( n \in \mathbb{N} \) and \( I_1, I_2 \) the canonical \( \delta_1 \)- and \( \delta_2 \)-interval trees, respectively. Then we have \( I_2^\delta \leq 7^{-|w|} \) for any \( w \in \mathbb{N}^{\delta_2} \). Define a computable function \( \iota : \mathbb{N}^{\delta_1} \to \mathbb{N}^\ast \) by \( \iota(\lambda) := \lambda \) and \( \iota(wa) := \iota(w)(6 - a) \), if \( a \in \{1, 3, 5\} \), and \( \iota(wa) := \uparrow \), otherwise. Obviously, \( \iota \) is \( (\delta_1, \delta_2) \)-compatible and effectively modular. By Theorem 4 there is a computable function \( f : [0; 1] \to [0; 1] \) of finite length such that \( f(I_1(w)) \subseteq I_2(\iota(w)) \) for any \( w \in \text{dom}(\iota) \).

Let \( (\varphi_i)_{i \in \mathbb{N}} \) be an effective enumeration of computable \( \varphi_i \subseteq \mathbb{N} \to \mathbb{Q} \). Then \( H := \{(i, n) : \varphi_i(n) \downarrow \text{ and } (\forall m < n (\varphi_i(m) < \varphi_i(m + 1))) \text{ or } (\forall m < n (\varphi_i(m) > \varphi_i(m + 1)))\} \) is an r.e. set. Especially, there is an effective enumeration function \( g : \mathbb{N} \to H \) such that

\[
g(s) = (i, n) \& n \neq 0 \implies \exists s' < s(g(s') = (i, n - 1)). \tag{1}
\]

We construct now a computable increasing sequence \( (x_s)_{s \in \mathbb{N}} \) of rational numbers by \( x_s := a_{w_s}^\delta \) for some computable sequence \( (w_s)_{s \in \mathbb{N}} \) of \( \text{dom}(I_1) \) which is constructed in stages as follows.

Stage 0: Define simply \( w_0 := “1” \).

Stage 1: Given \( w_s \in \text{dom}(I_1) \). Suppose that \( g(s) = (i, n) \). Define \( w_{s+1} := w_s 1 \) if \( \varphi_i(n) \notin I_2(\iota(w_s[i])) \) and, otherwise, define \( w_{s+1} \) of length \( \max\{i, |w_s|\} + 1 \) by

\[
w_{s+1}[m] := \begin{cases} (w_s[i][i] + 2 & \text{if } m = i \& \varphi_i(n - 1) > \varphi_i(n) \\ (w_s[i][i] + 4 & \text{if } m = i \& \varphi_i(n - 1) < \varphi_i(n) \\ (w_s[i]) & \text{if } m < i \\ 1 & \text{otherwise.} \end{cases}
\tag{2}
\]

We can show that \( (x_s)_{s \in \mathbb{N}} \) is a well defined computable increasing sequence of rational numbers which converges to \( x_\omega \) such that \( f(x_\omega) \) is not semi-computable.

Next Theorem shows that the class \( C_{wc} \) is not closed under monotone total computable real functions which have, of course, finite length.

**Theorem 6.** There is a computable increasing function \( f : [0; 1] \to [0; 1] \) and a weakly computable real number \( x \in [0; 1] \) such that \( f(x) \) is not weakly computable. Hence the class \( C_{wc} \) is not closed under \( \text{FIN} \) and \( C_{wc} \subseteq \text{FIN}(C_{wc}) \).

**Proof.** We define at first a recursive function \( \theta : \mathbb{N} \to \mathbb{N} \) inductively by \( \theta(0) := 5 \) and \( \theta(n + 1) := (\sum_{m \leq n} \theta(m) + 1) \cdot (n + 1) \cdot 5^{n+1} \). Furthermore, let \( \delta_1(n) := \theta(n) \) and \( \delta_2(n) := 5 \) be two recursive functions, \( I_1, I_2 \) the corresponding canonical \( \delta_1 \)-
and $\delta_2$-interval trees, respectively. Then we have $t_{w_s}^{\delta_2} \leq 2^{-|w|}$ at first. Define a computable function $\iota : \mathbb{N}_+^{\ast} \rightarrow \mathbb{N}_+^{\ast}$ with $\text{dom}(\iota) = \{1, 3\}^\ast$ by $\iota(w) := w$. Obviously, $\iota$ is also $(\delta_1, \delta_2)$-compatible and $(\delta_1, \delta_2)$-increasing. By Theorem 3 there is an increasing computable function $f : [0; 1] \rightarrow [0; 1]$ such that $f(\iota_1(w)) \subseteq I_2(\iota(w)) = I_2(w)$. Now we will construct a computable sequence $(x_s)_{s \in \mathbb{N}}$ of rational numbers which converges weakly effectively to some $x$ such that $f(x)$ is not weakly computable. Namely, $x$ satisfies, for all $i \in \mathbb{N}$, the following requirements:

$$R_i : \text{dom}(\varphi_i) = \mathbb{N} \& \sum_{n=0}^{\infty} |\varphi_i(n) - \varphi_i(n + 1)| < i \implies \forall n (|f(x) - \varphi_i(n)| \geq 5^{-i})$$

where $(\varphi_i)_{i \in \mathbb{N}}$ is an effective enumeration of computable functions $\varphi_i : \mathbb{N} \rightarrow \mathbb{Q}$. Let $H := \{\langle i, n \rangle : \forall m \leq n (\varphi_i(m) \downarrow)\}$. Then $H$ is r.e. and hence there is a recursive function $g : \mathbb{N} \rightarrow \mathbb{N}$ such that $\text{rang}(g) = H$. Especially, we can assume that $g$ satisfies condition (I) in the proof of Theorem 5. The sequence $(x_s)_{s \in \mathbb{N}}$ is defined by $x_s := \omega^{\delta_1}_{w_s}$ for the computable sequence $(w_s)_{s \in \mathbb{N}}$ of $\mathbb{N}_+^{\ast}$ which is constructed effectively in stages as follows.

Stage $s = 0$: Define $w_0 := \langle 1 \rangle$.

Stage $s+1$: Given $w_s \in \mathbb{N}_+^{\ast}$. Let $g(s) = \langle i, n \rangle$. If $\sum_{m<n} |\varphi_i(m) - \varphi_i(m + 1)| \geq i$ or $\varphi_i(n) \not\in I_2(\iota(w_s[i]))$, then define simply $w_{s+1} := w_s$ and hence $x_{s+1} := x_s$. Otherwise we define $w_{s+1}$ of length $i + 1$ by

$$w_{s+1}[m] := \begin{cases} 1 & \text{if } m = i \& w_s[m] = 3 \\ 3 & \text{if } m = i \& w_s[m] = 1 \\ w_s[m] & \text{if } m < i \\ \uparrow & \text{otherwise.} \end{cases}$$

(3)

In this case, we call the stage $s + 1$ an $i$-stage.

We can show that there are at most $i \cdot 5^i$ $i$-stages between stages $s_1$ and $s$, for any $i$ and $s_1 < s$ such that $s_1$ is an $i_1$-stage for some $i_1 < i$ or $s_1 = 0$ and there is no $j$-stage between $s_1$ and $s$ for any $j < i$. Therefore, there are at most $\theta(i)$ $i$-stages totally. Since $|x_s - x_{s+1}| = \prod_{j \leq i} \delta_1^{-1} \leq \delta_1^{-1}$, the sequence $(x_s)_{s \in \mathbb{N}}$ satisfies $\sum_{n=0}^{\infty} |x_s - x_{s+1}| \leq 2$. Namely, $x := \lim_{s \rightarrow \infty} x_s$ exists and is w.c. On the other hand, we can also show that $R_i$ is satisfied for any $i$ and hence $f(x)$ is not weakly computable.

**Corollary 1.** $\text{FIN}(C_{sc}) \subsetneq \text{FIN}(C_{wc})$.

## 5 Closure Properties under TOT

Different from the case for $\text{FIN}(C_{sc})$ and $\text{FIN}(C_{wc})$ we will show in this section that $\text{TOT}(C_{sc})$ and $\text{TOT}(C_{wc})$ coincide. Furthermore $\text{TOT}(C_{wc})$ is an algebraic field which is strictly between $C_{wc}$ and $C_{ra}$. This class can also be described by the limits of computable sequences of rational numbers with some kind of restriction on their convergence. Since $\text{TOT}^*(C) = \text{TOT}(C)$ for any $C \subseteq \mathbb{R}$ it suffices to mention only the results about $\text{TOT}(C)$.

At first, we introduce a kind of pairing function of real numbers.
Definition 3. The pairing function $\oplus : \mathbb{R}^2 \rightarrow \mathbb{R}$ is defined, for infinite $A, B \subseteq \mathbb{N}$ by

$$x_A \oplus x_B := \sum \{5^{-(2i+1)} : i \in A \} \cup \{3 \cdot 5^{-(2i+1)} : i \in A \} \cup \{5^{-(2i+2)} : i \in B \} \cup \{3 \cdot 5^{-(2i+2)} : i \notin B \}.$$ 

The pairing function $\oplus$ is unfortunately not computable. However its two inverse functions are computable as shown in the following lemma.

Lemma 2. The pairing function $\oplus : \mathbb{R}^2 \rightarrow \mathbb{R}$ satisfies the following conditions:

1. If $x$ and $y$ are left computable, then $x \oplus y$ is also left computable;
2. There are computable functions $\pi_1, \pi_2 : [0; 1]^2 \rightarrow [0; 1]$ such that, for any $x, y \in [0; 1], \pi_1(x \oplus y) = x$, $\pi_2(x \oplus y) = y$ and $\pi_1(x) \oplus \pi_2(x) = x$.

Theorem 7. 1. $C_{wc} \subseteq \text{TOT}(C_{lc});$
2. $\text{TOT}(C_{wc}) = \text{TOT}(C_{sc})$ and $\text{TOT}(C_{wc})$ is an algebraic field.

Now we show that the closure $\text{TOT}(C_{wc})$ does not exhaust the class $C_{ra}$.

Theorem 8. $C_{wc} \not\subseteq \text{TOT}(C_{wc}) \subset C_{ra}$

Proof. We sketch only the idea for $\text{TOT}(C_{wc}) = \text{TOT}(C_{lc}) \subset C_{ra}$. Let $(\alpha_n)_{n \in \mathbb{N}}, (\beta_n)_{n \in \mathbb{N}}$ and $(\gamma_n)_{n \in \mathbb{N}}$ be effective enumerations of computable functions $\alpha_n : \mathbb{N} \rightarrow \mathbb{N}$, $\beta_n : \mathbb{N} \times \mathbb{Q} \rightarrow \mathbb{Q}$ and $\gamma_n : \mathbb{N} \rightarrow \mathbb{Q}$, respectively. For any $a \in \text{TOT}(C_{lc})$, there is a computable function $f : [0; 1] \rightarrow \mathbb{R}$ and a computable increasing sequence $(\gamma_k(n))_{n \in \mathbb{N}}$ of rational numbers such that $b := \lim_{n \rightarrow \infty} \gamma_k(n) \in [0; 1]$ and $f(b) = a$. By Lemma 4, $f$ has a computable determinant function $\beta_j : \mathbb{N} \times \mathbb{Q} \rightarrow \mathbb{Q}$. Namely, $|\beta_j(n, r) - f(r)| \leq 2^{-n}$ for any $n \in \mathbb{N}$ and $r \in [0; 1]$. Furthermore, by Theorem 2.3, $f$ has a computable modulus $\alpha_i : \mathbb{N} \rightarrow \mathbb{N}$ such that for any $x, y \in [0; 1] \forall n \in \mathbb{N}(|x - y| \leq 2^{-\alpha_i(n)} \Rightarrow |f(x) - f(y)| < 6^{-n})$.

Therefore, it suffices to construct a computable sequence $(x_s)_{s \in \mathbb{N}}$ of rational numbers such that its limit $x := \lim_{s \rightarrow \infty} x_s$ exists and satisfies, for all $e := (i, j, k) \in \mathbb{N}$, the following requirements:

$$R_e : \text{If } \alpha_i, \beta_j, \gamma_k \text{ are total and } (\gamma_k(n))_{n \in \mathbb{N}} \text{ converges increasingly to } \gamma_k \text{ then } \exists m(|\gamma_k(m) - \gamma_k| \leq 2^{-\alpha_i(m)} \& |\beta_j(e, \gamma_k(m)) - x| \geq 6^{-(e+1)}).$$

To satisfy $R_e (e := (i, j, k))$, we choose a rational interval $(a; b)$ of length $6^{-e}$ as a base interval. This interval is divided by $a := a_0 < a_1 < a_2 < a_3 < a_4 < a_5 < a_6 := b$ into six subintervals $I_t$ for $t < 6$ of the same length $6^{-(e+1)}$. We wait for some stage $s$ such that $\alpha_i(e) := \alpha_i(s(e))$ is defined. Then look for the largest $m \leq s$ such that $\gamma_k,s(0) \leq \gamma_k,s(1) \leq \cdots \leq \gamma_k,s(m) \leq 1$ are all defined. If $\beta_j,s(e, \gamma_k,s(m))$ is also defined and located in the interval $I_t$ for some $t < 6$, then define $x := a_1$, if $t \geq 3$ and $x := a_4$; otherwise, if there is no such $s$, then define simply $x := a_1$. If at a later stage $s_1 > s$, we find some $m_1 > m$ such that $|\gamma_k,s_1(m_1) - \gamma_k,s_1(m)| \geq 2^{-\alpha_i(e)}$, then we repeat the above procedure for
Weakly Computable Real Numbers and Total Computable Real Functions

$m_1$ again, and so on. This can happen only finitely often, if $\gamma_k(n)$ converges to some $\gamma_k \in [0; 1]$ increasingly. Thus we can always successfully find such $x$ which satisfies $R_x$.

To satisfy all requirements simultaneously, we apply the above strategies for $R_x$ to an interval $I(w)$ with $|w| = e$ for the canonical $\delta$-interval tree $I$ on $[0; 1]$ for $\delta := 6$ by the priority injury technique. The details are omitted here.

At last we characterize the class $\text{TOT}(\mathbf{C}_{wc})$ by some convergence property of computable sequences of rational numbers.

**Definition 4.** Let $(a_i)_{i \in \mathbb{N}}$ be any sequence of real numbers and $n \in \mathbb{N}$. We define the degree of $n$-divergence of the sequence $(a_i)_{i \in \mathbb{N}}$ as the maximal natural number $m$ which satisfies the condition that there is a chain of natural numbers $i_0 < j_0 \leq i_1 < j_1 \leq \cdots \leq i_m < j_m$ such that $|a_{i_s} - a_{j_s}| \geq 2^{-n}$ for any $s \leq m$.

**Theorem 9.** The real number $x$ belongs to $\text{TOT}(\mathbf{C}_{wc})$ iff there is a computable sequence $(a_i)_{i \in \mathbb{N}}$ of rational numbers which converges to $x$ and a recursive function $f : \mathbb{N} \to \mathbb{N}$ such that, for all $n \in \mathbb{N}$, $\alpha(n) \leq f(n)$, where $\alpha(n)$ is the degree of $n$-divergence of $(a_i)_{i \in \mathbb{N}}$.

References

Turing Computability of a Nonlinear Schrödinger Propagator

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Abstract. We study Turing computability of the nonlinear solution operator \( S \) of the Cauchy problem for the Schrödinger equation of the form

\[
\frac{du}{dt} = -\frac{d^2 u}{dx^2} + mu + |u|^2 u
\]

in \( \mathbb{R} \). We prove that \( S \) is a computable operator from \( H^1(\mathbb{R}) \) to \( C(\mathbb{R}; H^1(\mathbb{R})) \) with respect to the canonical representations.

There is a rich mathematical theory for the Schrödinger equation of the form

\[
\frac{du}{dt} = (-\triangle + m)u + |u|^2 u, \quad t \in \mathbb{R}, \; x \in \mathbb{R}^n,
\]

where \( \triangle \) is the Laplace operator in \( \mathbb{R}^n \) and \( m \) is a real constant. The equation is widely used in several domains of applied physics. For example, it can be considered as the classical approximation to the field equation for a quantum mechanical nonrelativistic many body system with a two body \( \delta \)-function interaction. It is also used in the Landau-Ginsburg theory of superconductivity.

Our interest is to study computability of the solution operator. It is commonly believed among physicists that for processes which can be described by well-established theories, such as quantum systems described by the Schrödinger equation, the future behavior can be computed with arbitrary precision at least in principle from sufficiently precisely given initial conditions, where the computations can be performed on digital computers, hence on Turing machines. In this note we put the common belief into a formal basis of computable analysis, which is the theory of real number computation. We study computability of the nonlinear Schrödinger propagator within the framework of Type-2 Theory of Effectivity (TTE) \cite{Wei00}, which is based on work by Turing, Grzegorczyk, Lacombe, Ko, Kreitz and Weihrauch, Pour-El and Richards, and many others \cite{Tur36, Grz57, KW85, PER89, Ko91}. For simplicity of notation we consider only the case of space dimension \( n = 1 \).

The main result is the following theorem.
Theorem 1 The solution operator $S : H^1(\mathbb{R}) \to C(\mathbb{R}; H^1(\mathbb{R}))$ of the Cauchy problem

$$i \frac{du}{dt} = (-\Delta + m)u + |u|^2u, \quad t \in \mathbb{R}, \quad x \in \mathbb{R},$$

(1)

$$u(0)(x) = \varphi(x), \quad \varphi \in H^1(\mathbb{R}),$$

(2)

where $S(\varphi) = u$ and $m$ is a computable real number, is $(\delta_{H^1}, [\rho \to \delta_{H^1}])$-computable.

We use the framework of TTE [Wei00]. Let $L^2(\mathbb{R})$ denote the set of all $L^2$-functions defined on $\mathbb{R}$ equipped with $L^2$-norm, $H^s(\mathbb{R})$ the Sobolev space, $s \geq 0$, $C(\mathbb{R}, H^s(\mathbb{R}))$ the set of all continuous functions $g : \mathbb{R} \to H^s(\mathbb{R})$ equipped with the compact-open topology, and $S(\mathbb{R})$ the Schwartz space defined by

$$S(\mathbb{R}) = \{ \phi \in C^\infty(\mathbb{R}) : \forall \alpha, \beta \in \mathbb{N}, \sup_{x \in \mathbb{R}} |x^\alpha \phi^{(\beta)}(x)| < \infty \}.$$

Let $\delta_{L^2}$ and $\delta_{H^s}$ be the standard representations of $L^2(\mathbb{R})$ and $H^s(\mathbb{R})$, respectively, $[\rho \to \delta_{H^s}]$ the representation of $C(\mathbb{R}, H^s(\mathbb{R}))$, and $\delta_{SC}$ the Cauchy representation of $S(\mathbb{R})$. For the details about these representations the reader is refereed to [Wei00], [WZ99], and [WZ01]. We need another representation, denoted as $\tilde{\delta}_{H^s}$, of $H^s(\mathbb{R})$ [WZ01]. The representation $\tilde{\delta}_{H^s}$ makes use of the classical result that the set of Schwartz functions is dense in $H^s(\mathbb{R})$.

Lemma 2 The representation $\delta_{H^s}$ is equivalent to the representation $\tilde{\delta}_{H^s}$, if $s \in \mathbb{R}$ is computable.

Proof: See [WZ01].

We now consider the Cauchy problem [12]. We use a fixed point argument to construct a Type-2 Turing machine which computes a $[\rho \to \delta_{H^1}]$-name of the solution when input a $\delta_{H^1}$-name of the initial function. The Cauchy problem (1, 2) can be recasted in the form of the integral equation

$$u(t) = U(t - t_0, \psi) - i \int_{t_0}^{t} U(t - \tau, |u(\tau)|^2 u(\tau))d\tau$$

(3)

with $u(t_0) = \psi$. We will first construct a segment of $u$ on the interval $[t_0 - T_\varphi, t_0 + T_\varphi]$ by a fixed point iteration, where $T_\varphi$ is a constant depending only on the initial condition $\varphi$ of [12]. The following two conservation laws are needed.

Lemma 3 If $u : \mathbb{R} \to H^1(\mathbb{R})$ is a solution of [1], then for any $t, t', t_0 \in \mathbb{R}$,

$$||u(t)||_{H^1} \leq f(||u(t_0)||_{H^1})$$

(4)

$$||U(t - t', u(t'))||_{H^1} \leq f(||u(t_0)||_{H^1}),$$

(5)

where $f(x) = x \cdot (1 + 4\sqrt{x})^{1/2}$ and $U : \mathbb{R} \times S(\mathbb{R}) \to S(\mathbb{R})$ is the free evolution defined by $U(t, \psi) := (2\pi)^{-1/2} \int_\mathbb{R} e^{i\xi x - i|\xi|^2 t - i\xi t} \hat{\psi} d\xi$. 


Proof: See [WZ01].

We define two maps $A$ and $G$. The map $A : \mathbb{R} \times H^1(\mathbb{R}) \times C(\mathbb{R}, H^1(\mathbb{R})) \to C(\mathbb{R}, H^1(\mathbb{R}))$ is defined by

$$A(t_0, \psi)(v)(t) = U(t - t_0, \psi) - i \int_{t_0}^{t} U(t - \tau, |v(\tau)|^2 v(\tau))d\tau,$$

and the map $G : \mathbb{R} \times C(\mathbb{R}, H^1(\mathbb{R})) \to C(\mathbb{R}, H^1(\mathbb{R}))$ is defined by

$$G(t_0, v)(t) = -i \int_{t_0}^{t} U(t - \tau, |v(\tau)|^2 v(\tau))d\tau.$$

On a small neighborhood $I$ of $t_0$, $A(t_0, \psi)$ is a contraction on $C(I, H^1(\mathbb{R}))$, if $\psi$ is not too big.

**Lemma 4** Let $r_\varphi$ be a rational number such that $r_\varphi - 1 \leq ||\varphi||_{H^1} \leq r_\varphi$, let $R_\varphi$ be the least integer upper bound of $f(r_\varphi) + 1$ and let $I := [t_0 - T_\varphi, t_0 + T_\varphi]$, where $T_\varphi = 1/(32R_\varphi^2)$. Then for any $\psi \in H^1(\mathbb{R})$ such that $||\psi||_{H^1} \leq R_\varphi$ and any $v_1, v_2 \in C(\mathbb{R}; H^1(\mathbb{R}))$,

$$||A(t_0, \psi)v_1 - A(t_0, \psi)v_2||_I = ||G(t_0, v_1) - G(t_0, v_2)||_I \leq \frac{1}{2}||v_1 - v_2||_I \quad (6)$$

where $||v||_I := ||v||_{C(I, H^1(\mathbb{R}))} = \sup_{t \in I} ||v(t)||_{H^1(\mathbb{R})}$.

Proof: See [GV79].

The fixed point of the contraction $A(t_0, \psi)$ is the solution of $u(t_0) = \psi$. We are able to show that the restriction of the operator $A$ to $S(\mathbb{R})$ is computable (see Lemma 5 below). So we will approximate $\psi \in H^1(\mathbb{R})$ by a Cauchy sequence $\psi_k$ of $S(\mathbb{R})$-functions (see Lemma 2) and use approximations of the fixed points of $A(t_0, \psi_k)$ to compute the fixed point of $A(t_0, \psi)$ (see Lemma 6 below).

**Lemma 5** The restriction $B$ of the operator $A$ to $S(\mathbb{R})$ is

$$(\rho, \delta_{SC}, [\rho \to \delta_{SC}], [\rho \to \delta_{SC}])$$-computable.

Proof: See [WZ01].

**Lemma 6** Let $||\psi||_{H^1} \leq R_\varphi$ and $||\psi - \psi_k||_{H^1} \leq 2^{-k}$ for $k = 0, 1, \ldots$, where $\psi_k$'s are $S(\mathbb{R})$-functions. Let $v_\psi$ be the fixed point of $A(t_0, \psi)$. Then there are computable functions $g_1, g_2 : \mathbb{N} \times \mathbb{N} \to \mathbb{N}$ such that

$$||v_\psi - (A(t_0, \psi_{g_2(R_\varphi,m)}))^{g_1(R_\varphi,m)}(0)||_I \leq 2^{-m}$$
Proof: See [WZ01].

Combining lemmas 4–6 we can construct a Type-Two Turing machine which computes $u(t)$ from the initial condition $\varphi \in H^1(\mathbb{R})$ and arbitrary $t \in \mathbb{R}$.

Proof of Theorem 1: See [WZ01].

References


Author Index

Ahn, Hee-Kap 237
Aichholzer, Oswin 101
Alber, Jochen 318
Arslan, Abdullah N. 111
Aurenhammer, Franz 101
Bach, Eric 473
Bespamyatnikh, Sergei 121
Bläser, Markus 1
Bose, Prosenjit 142
Braumühl, Burchard von 586
Caballero-Gil, Pino 257
Cai, Jin-Yi 473
Cai, Maocheng 247
Castanho, Carla Denise 59
Chan, Wun-Tat 432
Chandran, L. Sunil 308, 384
Chang, Ee-Chien 453
Chen, Danny Z. 59
Chen, Guangting 519
Chen, Qingfeng 561
Chen, Wei 59
Chen, Zhixiang 121, 540
Cheng, Siu-Wing 237
Cheong, Otfried 159, 237
Chrobak, Marek 483
Chwa, Kyung-Yong 443
Cohen, Myra B. 420
Colbourn, Charles J. 420
Dai, H.K. 228
Dang, Zhe 529
Dawande, Milind 409
Deng, Xiaotie 247
Díaz, Josep 298
Du, Ding-Zhu 375, 509
Duncan, Rob 191
Eğecioğlu, Ömer 111
Epstein, Leah 463
Fang, Qizhi 247
Feng, Haodi 414
Fernau, Henning 318
Finocchi, Irene 357, 368
Froncek, Dalibor 420
Fujiwara, Akihiro 59
Gąsieniec, Leszek 483
Gengler, Romain 586
Golin, Mordecai 237
Gu, Xun 86
Heath, Samuel A. 64
Hemaspaandra, Lane A. 49
Hempel, Harald 49
Hernández-Goya, Candelaria 257
Hsu, Wen-Lian 207
Hu, Xiaobo (Sharon) 150
Hurtado, Ferran 101
Ibarra, Oscar H. 529
Igarashi, Yoshihide 571
Iwama, Kazuo 28
Kang, Liying 375
Kemmerer, Richard A. 529
Kikuchi, Yosuke 400
Kim, Hyun-Sung 262
Kim, Jae-Hoon 443
King, Valerie 268
Koucký, Michal 11
Krasser, Hannes 101
Lam, Tak-Wah 432
Li, Xiang-Yang 91, 181
Lou, Dingjun 349
Lu, Chin Lung 338
Luo, Xudong 561
Madhavan, C.E. Veni 328
Morin, Pat 142
Motegi, Kazuhiro 571
Motoki, Mitsuo 502
Nakano, Shin-ichi 131
Narayanaswamy, N.S. 328
Nardelli, Enrico 278
Nickelsen, Arfst 49
Niedermeier, Rolf 318
Nishizeki, Takao 288
Obokata, Kumiko  571
Okabe, Yasuo  28
Okamoto, Kyoichi  493
Omori, Michiko  571
Oostrum, René van  237

Petersen, Holger  218
Petreschi, Rossella  368
Preparata, Franco P.  64
Proietti, Guido  278

Qian, Jianbo  191
Qiao, Hong  375

Rettinger, Robert  586
Rytter, Wojciech  483

Serna, Maria  298
Shan, Erfang  375
Shibata, Yukio  400
Shin, Chan-Su  159
Siebert, Bodo  75
Stewart, Iain A.  39
Su, Kaile  561
Su, Zheng-Yao  338
Subramani, K.  550
Sung, Ting-Yi  390

Takase, Toshiro  28
Tan, Xuehou  170, 201
Tang, Chuan Yi  338
Thilikos, Dimitrios M.  298
Thorup, Mikkel  268

Ting, Hing-Fung  432
Tran, Nicholas  21

Vigneron, Antoine  142, 159

Wada, Koichi  59
Wang, Huaiqing  561
Wang, Jeng-Jung  390
Wang, Kanliang  121
Wang, Lusheng  509
Wang, Yu  181
Watanabe, Osamu  493
Weihrauch, Klaus  596
Widmayer, Peter  278
Wong, Wai-Ha  432
Wu, Guohua  576
Wu, Shiquan  86
Wu, Xiaodong  150

Xu, Baogang  395, 509
Xue, Guoliang  519

Yap, Chee  453
Yoo, Kee-Young  262

Zhang, Chengqi  561
Zhang, Shichao  561
Zheng, Xizhong  586
Zhong, Ning (Cincinnati)  596
Zhong, Ning (Guangzhou)  349
Zhou, Xiao  288
Zhu, Binhai  121, 191
Zhu, Shanfeng  247