A New Routing Scheme Based on Adaptive Selection of Geographic Directions

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Abstract—Geographic routing is recognized as an appealing approach to achieve efficient communications with low computational complexity and space cost. In order to apply this technology in Cyber-Physical Systems (CPSs), a comprehensive consideration must be given to performance issues such as throughput, delay, and load balance. In this paper, we provide a new routing scheme based on forwarding packets to multiple geographic directions. The proposed routing protocols are studied and analyzed theoretically. Theoretical bounds of throughput, delays and space cost are presented. Simulations show that our method performs more efficiently than traditional geographic routing schemes in terms of throughput, delay, and load balance with acceptable space cost. Our experiments also verify the tradeoff between performance metrics.

Index Terms—Geographic routing, protocol, cyber-physical systems, multi-direction routing

I. INTRODUCTION

With the emergence of Cyber-Physical Systems (CPSs) [17], wired or wireless network becomes an infrastructure to connect embedded and mobile devices. At the same time, the rapid development of localization technologies for indoor or outdoor environments makes it possible to obtain geographic information from mobile devices. Geographic routing, which utilizes physical locations of entities to deliver messages, becomes a promising way to provide robustness, energy saving and low cost for large-scale CPS networks [12, 21].

In most cases, applications in CPSs ask for good performance in terms of throughput, end-to-end delay, load balance, and guaranteed delivery of messages. These issues were not fully studied in previous research of geographic routing. In addition, most of the existing work on geographic routing focuses on only one or two of these issues. However, these performance issues must be considered comprehensively in order to meet rigid requirements of CPS network. In this paper, we provide a new geographic routing scheme that takes into account all of the above performance issues. The purpose of our method is to use acceptable space cost to achieve good performance.

Our strategy is to provide a new geographic routing scheme, which will forward packets to multiple geographic directions (Multi-Direction Routing, MDR for short). After assigning multiple directions to a network, nodes are sorted into global sequences in each direction. A delivery guaranteed protocol is proposed to send/receive messages based on the idea of virtual links. Based on this protocol, an efficient adaptive forwarding method called minimum angle policy is presented. The difference between MDR protocol and other existing geographic routing protocols is presented in Section 2.

We apply theoretical bounds on throughput, end-to-end delay, stability and the space complexity of the proposed routing protocol. From our analysis, we prove that throughput that is achieved by our method is polynomially proportional to the number of directions assigned to a network, while end-to-end delay that is achieved is the opposite. Also, space complexity of the proposed protocol is the same as that of the classical Shortest Path Routing (SPR) protocol.

When we compare the performance of our approach with several typical routing schemes, the experimental results show that our method performs much better in terms of throughput, end-to-end delay and load balance. Moreover, our method can achieve good performance with acceptable space cost. In addition, we provide a classification for geographic routing schemes that reveal cost-performance tradeoff in different routing methods. That is, higher space cost may bring more improvements on performance and vice versa. This insight will be helpful for choosing a proper routing

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scheme when various limitations in CPS applications are considered.

Our study has been organized as follows: Section II gives the related work including a classification on existing geographic routing schemes; Section III presents the details of the routing scheme proposed; Section IV offers theoretical analyses on throughput, end-to-end delay, stability and space cost; Section V compares our approach with other classical routing methods; and our conclusion is in Section VI.

II. RELATED WORK

We provide a classification of existing geographic routing schemes to investigate merits of our proposed method. Basically, geographic routing schemes can be classified into two types: protocols with space cost and protocols without space cost. The former type needs to record route information. The later one only needs to know location information of neighbors of a node. On the other hand, according to their forwarding policies, geographic routing schemes can also be classified into adaptive forwarding and static forwarding. The former can dynamically change policies in the forwarding process; while the later one determines the policy once a message has been sent out. Based on the above analysis, we provide a classification in Table 2.1.

<table>
<thead>
<tr>
<th>Protocols</th>
<th>Forwarding methods</th>
<th>Adaptive</th>
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<tbody>
<tr>
<td>With space cost</td>
<td>I. [5, 7, 12, 14, 15, 16, 19, 22, 28]</td>
<td></td>
</tr>
<tr>
<td>No space cost</td>
<td>II. Our work</td>
<td></td>
</tr>
<tr>
<td></td>
<td>III. [2, 3, 6, 10, 11, 13, 21, 23, 29]</td>
<td>IV. [4, 24, 25, 27, 38]</td>
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The typical method of Class I is face routing [5, 7, 12, 14, 15, 16, 19, 22, 28], while Class III includes the work of graph embedding [2, 3, 6, 10, 11, 13, 21, 23, 29]. Both types adopt the idea of greedy forwarding. A well-known problem of this strategy is that packets may get stuck in a local minimum unfortunately when no neighbors are closer to the destination. To solve this problem, face routing planarizes the network graph and forward packets along a sequence of adjacent faces towards destinations. It is obvious that this approach needs additional storage cost [7]. Graph embedding will assign virtual coordinates to nodes to ensure greedy forwarding always success. By this means, it does not need to record any route information. Though both methods need no or little storage space, their performance suffers from low throughput or long end-to-end delays. Actually, the nature of face routing is to cut some links to achieve guaranteed deliveries, which will seriously decrease the bandwidth capacity. For graph embedding, inconsistency of virtual coordinates and physical locations may bring remarkable transmission delays.

The work of Class IV also adopts routing protocols with no space cost. Typical methods of this class include [4, 24, 25, 27, 38]. For example, in [25], randomized approximations to throughput optimal routing were studied. The approach has no significant loss in throughput. It also has good load-balancing properties without compromising latency. However, this method uses multiple copies of packets during transmissions to avoid any cost on routing tables. Apparently this will place a heavy burden on network traffic. The feature of zero space cost makes this approach unable to guarantee delivery of messages, which is a must-have function in many circumstances.

We provide a new geographic routing scheme, which belongs to Class II. Compared with the work of Class I and III, our method can achieve better throughput, end-to-end delay and load balance with acceptable space cost. The extra space cost also enables our method provide guaranteed delivery of messages, which is lacking in routing schemes in Class IV.

III. MULTI-DIRECTION ROUTING SCHEME

In this section, we will first introduce steps of our proposed routing scheme. Then the formal routing protocol is presented and correctness of this approach is proved.

Similar to geographic routing schemes in [12, 14, 21], our approach (called Multi-Direction Routing, MDR for short) also uses geographic positions of nodes to deliver packets. However, MDR has remarkable differences from previous geographic routing methods:

A. Assigning multiple directions to a network

Firstly, we can assign multiple directions to a network. The direction space of a node is split equally by each direction. The bisector of the angle between two neighbor directions is the dividing line for the splitted direction space. In Fig. 3.1, the dashed line between direction 1 and 2 is the dividing line. When a packet is forwarded in an orientation with the smallest angle to a special direction such as direction 1, it is said the packet is forwarded to direction 1.

A requirement of this assignment is that any direction’s inverse direction must be included. We call such a pair of directions as Inverse Direction Pair (IDP for short). In Fig. 3.1, the direction 2 and 6 must be assigned as an IDP at the same time. Therefore, assigning one IDP can achieve mutual communications for any
node pair since any node can send packets to its subsequent and front nodes by two inversed directional sequences (e.g. (3-1) and (3-2) in Section 3.2). If a 2D plane is divided equally by IDPs, we say these IDPs construct an even assignment. Fig. 3.1 shows a 2D plane splitted by 4 IDPs equally.

![Figure 3.1. An assignment with 4 IDPs or 8 directions.](Image)

**B. Global sorting on geographic directions**

The nodes are sorted into a sequence according to their coordinate projections on geographic directions. For example, in Fig. 3.2, there are seven nodes \( n_1, n_2, \ldots, n_6 \) and \( n_7 \) in a 2D plane. Denoting their projections on the direction \( j \) as \( s_1, s_2, \ldots, s_6 \) and \( s_7 \). According to the numerical order (from small to large) of projections, we get a sequence as follows (such a sequence is called directional sequence):

\[
n_1 < n_2 < n_3 < n_4 < n_5 < n_6 < n_7 \quad (3-1)
\]

Denoting the inverse direction of the direction \( j \) as \( \bar{j} \), we get a directional sequence for \( \bar{j} \):

\[
n_1 > n_2 > n_3 > n_4 > n_5 > n_6 > n_7 \quad (3-2)
\]

The direction \( j \) and its inverse direction \( \bar{j} \) compose an IDP.

If any two adjoined nodes in a directional sequence (e.g. \( n_3 \) and \( n_4 \), or \( n_5 \) and \( n_6 \) in Fig. 3.2) are directly linked, it is obvious that any node can send packets to any of its subsequent nodes if all nodes relay packets to their direct subsequent nodes. However, if the adjoined nodes are not directly linked (e.g. \( n_1 \) and \( n_2 \) in Fig. 3.2), we will construct a virtual link (i.e. finding a path between these two nodes) by which a packet can be delivered between the above two adjoined nodes. In fact, for \( n_1 \) and \( n_2 \), a feasible path does exist, that is, from \( n_1 \) to \( n_3 \), to \( n_4 \) and to \( n_2 \). Then this path is regarded, as a virtual link from \( n_1 \) to \( n_2 \) and it will be recorded in \( n_1, n_3 \) and \( n_4 \) respectively. According to whether two adjoined nodes are directly linked or not, a node will forward a packet along the virtual link if such a link exists; otherwise, the node will forward the packet to its direct subsequent node. In Fig. 3.2, \( n_3 \) records the virtual link from \( n_1 \) to \( n_2 \). When \( n_4 \) receives a packet whose destination is \( n_2 \), it forwards the packet to \( n_2 \) but not its subsequent node \( n_5 \). Nevertheless, for a packet from \( n_5 \) to \( n_7 \), \( n_6 \) will forward it to \( n_7 \) without the aid of virtual links since two adjoined node pair \( n_5 \) and \( n_6 \), \( n_6 \) and \( n_7 \) are directly linked.

![Figure 3.2. An example of Multi-Direction Routing](Image)

**Algorithm 3.1. Sorting Protocol**

**Input:** Physical positions of \( N \) nodes and \( M \) directions in local area of node \( n_i \)

**Output:** The routing table for node \( n_i \) in all directions (taken the \( j \) th direction as an example)
Adaptive forwarding by selecting directions

After constructing routing tables for all nodes, we need to design strategies to forward packets. As concerned, the routing algorithm should achieve good performance in capacity. The forwarding policy is designed based on it. In the MDR scheme, a node can calculate the direction constructed by the source, current location and destination of a packet. We call such a direction as target direction. Adaptive forwarding will be achieved if an optimal direction is dynamically selected.

Before we present the forwarding method, we need to explain the interference model first. In ad-hoc networks, transmission links with closely spaced location have mutual interference with each other. It forms a transfer constraint between links that only a portion of links can transmit simultaneously, while the other links of interference cannot be used. Typically, interference can be grouped into two types. The first one typically is called node-exclusive interference model [36]. It considers that any two links with a common node, that is 1-hop away, cannot transmit simultaneously. The other model mainly considers the MAC technique in IEEE 802.11 protocol. The nodes should exchange the handshake control message (RTS-CTS before transmitting and DATA-ACK when transmitting) between each other [37].

A distributed scheduling model is built based on these interference models. Regard the links that can interfere the transmitting of link as interference set \( I(l) \). Any interference model should discover the set \( I(l) \). This set is considered already known for general purpose in this paper, no matter which specific interference model is applied. The scheduling model in this paper consists of two parts: distributed link schedule that optimizes the flow control in all directions and distributed packet schedule that optimizes the flow dispatching in all directions.

1) Distributed maximal weight link schedule for all directions

This schedule gives the flow control policy over links for all directions. The schedule focuses on maximizing the throughput capacity as large as possible. For easy understanding, we explain distributed directional link schedule in detailed steps. Then, we will give the illustration and formal description.

Let \( l = n_i n_j \) represent the link that node \( n_i \) to node \( n_j \) in direction \( j \). We define \( Q_l \) as the queue length of link \( l \) in node \( n_i \), and \( C_l \) as the transmitting data rate capacity of link \( l \) from node \( n_i \) to node \( n_j \). The links in interference set \( I(l) \) and the end nodes of these links construct the local schedule graph \( G_{\text{local}}(n_i) \) for node \( n_i \). Let \( S(n_i) \) denote the state of node \( n_i \), in which,
\( S(n_i) \in S = \{ \text{uncertainty, transmission, non-transmission} \} \).

At the beginning, all nodes are at the state of uncertainty.

The distributed maximal weight link schedule follows these steps:

a) If node \( n_i \) is in uncertainty state, find the maximal weight direction of node \( n_i \), that is, get link \( l \) in this direction, \( l = n_i n_d \), in which \( n_d = \arg \max_{n_d \in \text{nb}or(n_i)} Q C_j \) and \( n_i = n_j \).

b) Find whether the link \( l = n_i n_d \) is the maximal weight link in its interference set \( l(l) \). If so, node \( n_i \) (or \( n_d \)) sets its state as transmission and informs this state change to its neighbor nodes. Else if node \( n_i \) was informed and \( n_i \neq n_d \), node \( n_i \) sets its state as non-transmission and informs this state change to its neighbor nodes. Otherwise, ignores the message.

c) If node \( n_i \) was informed the transmission message and \( n_i = n_d \), node \( n_i \) sets its state as transmission and informs this state change to its neighbor nodes.

d) If node \( n_i \) was informed the non-transmission state from node \( n_j \) and \( n_i \) is in uncertainty state, node \( n_i \) deletes node \( n_j \) and the links with \( n_j \) as end nodes in its local schedule graph \( G_{\text{local}}(n_i) \). Node \( n_i \) repeats the steps a) and b).

e) If node \( n_i \) (\( n_j = n_i \) or \( n_j = n_i' \)) finishes the transmission, node \( n_i \) sets its state as uncertainty and informs this state change to its neighbor nodes.

f) If node \( n_i \) is in non-transmission state and was informed a state change of transmission to uncertainty, node \( n_i \) sets its state to uncertainty and informs this state change to its neighbor nodes.

g) If node \( n_i \) was informed a state change of non-transmission to uncertainty from node \( n_j \) and has deleted node \( n_j \) and the links with \( n_j \) as end nodes in its local schedule graph \( G_{\text{local}}(n_i) \), node \( n_i \) re-adds node \( n_j \) and the links into graph \( G_{\text{local}}(n_i) \).

In summary, when a maximal weight link \( l \) is selected, the two end nodes of link \( l \) set their states to transmission. All neighbor nodes (1-hop away) of the two end nodes are informed and set their state to non-transmission. All neighbor nodes (2-hop away) of these 1-hop neighbor nodes are informed and delete non-transmission state nodes in their local schedule graphs \( G_{\text{local}} \). Finishing transmission, nodes in transmission and non-transmission go back to the state uncertainty and nodes re-add these deleted nodes and links in graph \( G_{\text{local}} \). In Fig. 3.3, we assume the links are duplex and \( l_{13} \) is the maximal link in \( I(l_{13}) \), in which \( I(l_{13}) = \{ l_{13}, l_{24}, l_{34}, l_{35}, l_{46}, l_{56} \} \). Let \( V(l_{13}) = \{ n_1, n_6, n_4, n_5 \} \). The local graph \( G_{\text{local}}(n_i) = V(l_{13}) \cup I(l_{13}) \), that is, the nodes within 2-hop and the links within 2-hop from the links such as \( l_{13} \) with \( n_1 \) as an end node. When link \( l_{13} \) transmits from \( n_1 \) to \( n_3 \), the node \( n_2 \) and \( n_6 \) should remove the nodes and links in \( G_{\text{local}}(n_i) \) from their local schedule graph. When link \( l_{13} \) finish transmitting, the node \( n_2 \) and \( n_6 \) re-add these deleted nodes and links.

![Figure 3.3. An example of link schedule](image)

The formal description is presented in Algorithm 3.2.

**Algorithm 3.2. Link Schedule**

**Input:** local schedule graph \( G_{\text{local}}(n_i) \), queue length \( Q \) and capacity \( C \) of each link for node \( n_i \)

**Output:** link \( l \) scheduled

**Begin**

1. Set \( L(n) = \{ l | l = mn'(nn') \} \)
2. Find the link \( l = n_i n_d \), \( n_d = \arg \max_{n_d \in \text{nb}or(n_i)} Q C_j \)
3. If \( l = \arg \max_{l \in L(n)} Q C_j \)
   
4. set state \( S(n_i) = \text{transmission} \), notify all neighbor nodes \( n, n \in \text{nb}or(n_i) \)
5. If neighbor node \( n = n_d \)
6. n set state \( S(n) = \text{transmission} \), notify all neighbor nodes \( n', n' \in \text{nb}or(n) \) and all \( n' \) set state \( S(n') = \text{non-transmission} \)
7. If neighbor node \( n \neq n_d \)
8. n set state \( S(n') = \text{non-transmission} \), notify all neighbor \( n', n' \in \text{nb}or(n) \) and all \( n' \) set its local schedule graph \( G_{\text{local}}(n') = G_{\text{local}}(n') - n - L(n) \)
9. When link \( l \) finish transmitting
10. n with \( S(n) = \text{transmission} \) or \( S(n) = \text{non-transmission} \) set its state \( S(n) = \text{uncertainty} \)
11. n with \( S(n) = \text{non-transmission} \) notify all neighbor nodes.
\[ n', n' \in \text{neighbor}(n) \text{ and all } n' \text{ set its local schedule} \]

\[ G_{\text{local}}(n') = G_{\text{local}}(n) + n + L(n) \]

\textbf{End}

2) \textit{Distributed minimum angle based packet schedule for all directions}

Distributed packet schedule optimizes the flow dispatching in all directions. We focus on an adaptive forwarding policy called \textit{minimum angle policy}. At the source node, it is easy to calculate the target direction since locations of the source and destination are known. The target direction will be compared with all directions assigned to the network. Then a closest direction (with the minimum angle) will be selected. After choosing a direction, the forwarding process is accomplished with the aid of routing tables. For example, in Fig. 3.2 or Fig. 3.3, when forwarding a packet from \( n_i \) to \( n_j \) in direction \( j \), we know that the nearest neighbor of \( n_i \) is \( n_j \). Since \( \text{rank}(n_i, j) \in [\text{rank}(n_i), j, \text{rank}(n_j)] \) exists, there is virtual links from \( n_i \) to \( n_j \). When forwarding packets from \( n_j \), we find the minimum angle to the final destination. That is, to find the vector \( \overrightarrow{d_0} \), nearest to which direction in set \( S \). If a direction \( j_k \) is with the minimum angle, node \( n_i \) checks the link \( l \) in direction \( j_k \).

b) If the queue length \( Q_l \geq T_l C_l \), remove direction \( j_k \) from \( S \). If \( S \neq \emptyset \), back to step a).

c) Get the items \( < j_k, \text{rank}(s_j, j_k) > \) and \( < j_k, \text{rank}(s_j, j_k) > \) from the item table of packet \( p \). If \( \text{rank}(s_j, j_k) > \text{rank}(n_i, j_k) \) or \( \text{rank}(s_j, j_k) > \text{rank}(n_i, j_k) \) or \( \text{rank}(n_i, j_k) \), remove direction \( j_k \) from \( S \). If \( S \), back to step a) (Judging if the condition is met in this step is the key to guarantee any packet can be delivered successfully, we will prove the correctness of the protocol based on it). Else overwrite \( < j_k, \text{rank}(n_i, j_k) > \) and \( < j_k, \text{rank}(n_i, j_k) > \) to the item table of packet \( p \), change the direction to \( j_k \) and overwrite the item \( 0 : < j_k, d > \) to \( < j_k, d > \).

d) Forward the packet \( p \) to the next node following the target direction recorded in item 0.

In summary, when a packet \( p \) reach node \( n_i \), the direction \( j \) has as possible minimum angle with vector \( \overrightarrow{d_0} \). The direction \( j \) must meet two conditions. The first one is the selected link \( l \) in direction \( j \) has the queue length \( Q_l \leq T_l C_l \). The other one is \( \text{rank}(n_i, j) > \text{rank}(s_j, j) \), in which \( \text{rank}(s_j, j) \) is
recorded in the item \( <j, \text{rank}(s_j, f)> \).

![Figure 3.4. An example of packet schedule](image)

In Fig. 3.4, a packet \( p \) is sent from node \( n_1 \) to node \( d_0 \). At first, the vector \( n_1 \hat{d}_0 \) has minimum angle with direction 1, so \( p \) is sent in this direction. When \( p \) reaches \( n_1 \), it alternates to direction 8 by the packet schedule. Along the routing track, \( p \) finally reaches destination, node \( d_0 \). The formal description is presented in Algorithm 3.3.

**Algorithm 3.3. Packet Schedule**

**Input:** Packet \( p \) (destination \( d_0 \)) and node \( n_j \)

**Output:** the direction for \( p \) delivery

**Begin**

1. If \( n_j = d_0 \), received
2. Set \( S = \{j | \text{rank}(d_0, f) > \text{rank}(n_j, f)\} \)
3. Set \( j_d = p(0, \text{dir}) \), in which \( p(0, \text{dir}) \) denotes selecting the direction of item 0 from packet \( p \) and \( p(j, \text{rank}) \) denotes selecting rank from item indexed with direction \( j \)
4. Set \( j_d' = j_d \)
5. While \( S \)
6. \( j_{\min} = \arg \min_{j\in S} \theta_j \), in which \( \theta_j \) denote the angle between direction \( j \) and \( n_1 \hat{d}_0 \) and select link \( l \) in \( j_{\min} \)
7. If \( Q_l \geq T_l C_l \) or \( p(j_{\min}, \text{rank}) > \text{rank}(n_j, j_{\min}) \) or \( p(j_0, \text{rank}) > \text{rank}(n_j, j_0) \) or \( \text{rank}(n_j, j_{\min}) > \text{rank}(d_0, j_{\min}) \)
8. \( S = S - j_{\min} \)
9. Else
10. Set \( p(j_{\min}, \text{rank}) = \text{rank}(n_j, j_{\min}), p(j_0, \text{rank}) = \text{rank}(n_j, j_0) \)
11. Set \( j_d' = j_{\min} \)
12. Break
13. If \( j_d \neq j_d' \)
14. Set \( p(0, \text{dir}) = j_d' \)
15. Deliver packet \( p \) along direction \( p(0, \text{dir}) \)

**End**

Note that in a distributed system, we do not actually know the global rank of nodes. The physical location of nodes could be record for rank comparison between nodes. Comparison like \( \text{rank}(s_j, f) > \text{rank}(n_j, j_k) \) actually is the comparison of perpendicular projection of the physical location of \( s_j \) and \( n_j \) in direction vector \( j_k \).

**D. Summary**

In the MDR protocol, node \( n_j \) firstly sorts nodes in directions and construct the routing table (Algorithm 3.1). Then, node \( n_j \) could use link schedule (Algorithm 3.2) to optimize the flow capacity in all directions and packet schedule (Algorithm 3.3) to smooth the flow in all directions. Link schedule and packet schedule could be carried out in parallel.

The contribution of the MDR protocol mainly lies in that it uses the direction as metric for routing, while current geographical routing protocol uses location as metric. As the metric changes from the location to the direction, the mostly concerned issue that avoiding packets be lost in local minimum needs a new solution. This is the reason why we develop Algorithm 3.1, 3.2 and 3.3. In Section 4 (Proposition 4.1), we will prove that packets with the MDR protocol can avoid local minimum to reach the destination. Also, in Section 4, we will give the theoretical analysis and show how these algorithms of the MDR protocol can improve the performance.

**IV. THEORETICAL ANALYSIS**

**A. The Model**

![Figure 4.1 illustrates the model we used to conduct theoretical results](image)

Figure 4.1 illustrates the model we used to conduct theoretical results. A random planar network is considered where \( N \) nodes are randomly distributed over a unit square whose size is 1x1. An even assignment with \( M \) directions or \( M/2 \) IDPs is used. The radio range of each node is \( r = R \sqrt{(\log N)/N} \), where \( R > 1/\sqrt{\pi} \) [8]. The region is further divided in tiles whose length and width are both \( r \) (i.e. there are total \( u = 1/r^2 \) tiles). There are \( N (\alpha \) is a scale factor and \( \alpha > 0 \) ) source-destination pairs randomly selected and the routing strategy makes at least \( \delta r \) (\( \delta \) is a constant factor and \( 0 < \delta < 1 \)) progress towards the destination in each forwarding step. Actually, \( \delta \) is inversely proportional to the total number of \( M \). When more directions are available (i.e. larger \( M \)), a better path closer to the target direction could be found, which means \( \delta \) can take a greater value in average.
B. The Correctness of The Protocol

First, we will prove the correctness of the MDR protocol that can guarantee delivery of packets.

**Proposition 4.1.** Any packet can be delivered successfully with the MDR protocol applied.

Proof: Firstly, we prove a packet delivered without alternating direction can be delivered successfully. Then, we explain an alternating direction does not affect the correctness of the MDP protocol.

Without loss of generality, we assume packet \( p \) is sent from \( n_i \) to \( n_k \) (here \( i \neq k \)) in direction \( j \) where \( \text{rank}(i, j) > \text{rank}(k, j) \). Let \( n_h \) be the node that the packet locates. At the beginning, we have \( n_h = n_i(h - k) \).

When the packet does not arrive at the destination, we can consider following three cases:

Case 1): if \( n_h = \text{nbort}(n_i) \), we can directly deliver this packet to node \( n_k \); set \( n_h = n_k \).

Case 2): if \( \text{rank}(h, j) > \text{rank}(k, j) \), \( Y \neq \emptyset \), where \( Y = \{n_l \mid \text{rank}(l, j) < \text{rank}(k, j), n_l = \text{nbort}(n_i)\} \), we know that there exists some neighbors in direction \( j \), and these nodes are prior to destination \( n_k \) in this direction. The packet is delivered to one of node \( n_l \) in \( Y \), and we set \( n_h = n_l \). Here the packet approaching the destination does not skip the destination \( n_k \) in this direction.

Case 3): if \( \text{rank}(h, j) < \text{rank}(k, j) \), \( Y \neq \emptyset \), we know that there must be virtual links that have been established. Then, along the virtual link, the packet is delivered to \( n_l \), which satisfies \( \text{rank}(h, j) < \text{rank}(l, j) < \text{rank}(k, j) \). Note that here \( n_l \) could be the destination node \( n_k \). Under this condition, the packet approaching the destination does not skip the destination \( n_k \) in this direction.

In conclusion, Case 1), 2) and 3) will not skip the destination \( n_k \), thus delivery without alternating direction \( j \) is guaranteed.

When alternating direction, packet \( p \) can also be delivered successfully. Assume packet \( p \) finally routes to destination through direction \( j' \). Let \( n_i \) represent the first node that \( p \) starts routing in direction \( j' \). Consider the following two cases:

Case 1): when packet \( p \) has arrived at node \( n_i \), it will be routed to the destination \( n_h \) without alternating direction. As proved above, packet \( p \) will be delivered successfully.

Case 2): when packet \( p \) has arrived at node \( n_i \), it will be routed to the destination \( n_h \) with alternating directions and finally back to \( j' \). As we look back the condition of alternating directions in Algorithm 3.3. When back to direction \( j' \), we have (1) the condition \( p(j', \text{rank}) > \text{rank}(n_i, j') \) must be met. Packet will not back to the nodes of lower rank that has passed by when back to direction \( j' \), so the MDR protocol can avoid the local minimum problem. (2) the condition \( \text{rank}(n_i, j_{\text{min}}) < \text{rank}(d_0, j_{\text{min}}) \) must be met. This means that the destination \( d_0 \) cannot be skipped in direction \( j' \).

As proved, any packet can be delivered successfully with the MDR protocol. ■

C. Analysis of Throughput

**Theorem 4.1.** If the MDR protocol and the minimum angle policy are adopted, every source-destination pair can achieve a data rate of \( \Theta(j / \sqrt{N \log N}) \) simultaneously.

Proof: first, we will prove that it is almost sure that the following inequality stands:

\[
\text{Hops} \leq \frac{4}{\delta} N \log N \quad (4-1)
\]

\[
H = \sqrt{2} \alpha R N \log N + 10 \sqrt{2} a R N \log N \quad (4-2)
\]

In (4-1), \( \text{Hops} \) indicates the total number of forwarding hops of all \( \alpha N \) source-destination pairs passing any single tile. For a specific tile, according to the MDR algorithm and the forwarding policy, some source-destination pairs may go through it via one or multiple hops. We use the random variable \( X_j \in \{0,1\} \) (\( 1 \leq i \leq 1/h^2 \), \( 1 \leq j \leq \alpha N \)) to indicate if the \( j \)th path will pass the \( i \)th tile. \( X_j = 1 \) means that the \( j \)th path will pass the \( i \)th tile, otherwise not. Then the total
number of source-destination pairs passing the $i$th tile will be $S_i = \sum_{j=1}^{a_i} X_i^j$. For the $j$th path, we know that the maximum number of its hops in the $i$th tile will be $h_i = \sqrt{2}/\delta$. Therefore, the total number of forwarding hops in the $i$th tile is upper bounded by the product of $h_i$ and the maximum of $S_i$. That is
\[
Hops \leq h \cdot \max(S_i) = \sqrt{2}/\delta \max(S_i)
\]  
\hspace{1cm} (4-3)

To obtain (4-1), we first prove that the following inequality has small probability of occurrence:
\[
P(Hops > \sqrt{2}/\delta - H) \text{ or } P(\max(S_i) > H) \hspace{1cm} (4-4)
\]

Here $P$ indicates the probability. From the knowledge of statistics, we know that the union bound of the probability is greater than the maximal of the probability. Then, we have
\[
P(\max(S_i) > H) \leq \sum_{i=1}^{\infty} P(S_i > H) \hspace{1cm} (4-5)
\]

Let $\tilde{X}_i^j$ be the Bernoulli random variable [20] stochastically dominating $X_i^j$. Then the probability of $P(\tilde{X}_i^j = 1)$ is greater than that of $P(X_i^j = 1)$. Denoting $\sum_{j=1}^{\infty} \tilde{X}_i^j$ as $\tilde{S}_i$, then we have
\[
P(S_i > H) = P(\tilde{S}_i > H) \hspace{1cm} (4-6)
\]

Putting (4-5) and (4-6) together, we have
\[
P(\max(S_i) > H) \leq \frac{1}{r^i} P(\tilde{S}_i > H) \hspace{1cm} (4-7)
\]

We instantiate $X_i^j$ and make it dominate $X_i^j$ surely and stochastically. Denoting $u_i$ as the upper bound of the number of tiles passed by the $j$th path. Since the number of jumps required to reach the destination for any route is no more than $\sqrt{2}/(r)$ hops ($\sqrt{2}$ is the maximal length between two nodes), we set $u_i = \sqrt{2}/(r\delta)$. So the probability that a tile may be touched by the $j$th path will be $P_\delta = u_i / u = \sqrt{2}/\delta$. Then, considering the random variable $P(\tilde{X}_i^j = 1) = P_\delta$. Since $P_\delta$ is the maximum probability of a tile touched by the $j$th path, $X_i^j$ surely stochastically dominates $X_i^j$ at the setting. So, we have $E(\tilde{X}_i^j) = \alpha N \cdot P_\delta$
\[
E(\tilde{X}_i^j) = \sqrt{2\alpha \delta R N \log N} \delta 
\]

Both $X_i^j$ and $X_i^j$ are Bernoulli random variables. $X_i^j$ is independent to $\tilde{X}_i^j$ if $j \neq k$ for any $i$ and $l$. For the sums of i.i.d. Bernoulli random variables, we have [1]
\[
P(\tilde{X}_i^j > H) \leq \frac{1}{N^3} \hspace{1cm} (4-9)
\]

Combining with (4-9), by summing over all $j$ in (4-7), we have
\[
P(\max(S_i) > H) \leq \frac{1}{R^2 N^4 \log N} \hspace{1cm} (4-10)
\]

Therefore $P(\max(S_i) > H)$ or $Hops > \sqrt{2}/\delta - H$ has small probability of occurrence. Thus, we prove that it is almost sure that (4-1) will converge.

Next, we will present a bound for throughput. The transmission of a node in a tile may affect transmissions of other nodes in the range influenced in current or neighboring tiles. We assume that at most $D$ neighbor tiles are affected. As refer to the graph coloring problem, all tiles can be colored by $D+1$ colors. Considering a time interval $T$ with the fixed length, we can divide $T$ into $T/(D+1)$ slots. Rewriting (4-1) as
\[
Hops \leq A \sqrt{\frac{N \log N}{N}} \hspace{1cm} (4-11)
\]
where $A$ is a finite number (note that $A \sim 1/\delta$ here). Thus, each forwarding hop in any tile can get a transmission time as
\[
\frac{1}{A(D+1)\sqrt{N \log N}} \hspace{1cm} (4-12)
\]

Therefore, we complete the proof that each source-destination pair can achieve the data rate of $(1/\sqrt{N \log N})$ simultaneously.

Note. Here $A \sim 1/\delta$ and $M \sim \delta$ means that when more directions are available, the MDR protocol has better throughput.

D. Analysis of End-to-end Delay

Theorem 4.2. If the MDR protocol and the adapt-
tive-based policy are adopted, the delay of any source-destination pair is no more than \( (N) \). Proof: we know that the total number of hops for a source-destination pair in the square is at most \( \sqrt{2} / \delta r \) (the maximal distance between two nodes is \( \sqrt{2} \), and the routing strategy progresses \( \delta r \) at least in each forwarding step). As proved in Theorem 4.1, each source-destination pair can achieve a data rate of \( \Theta(1/N \log N) \) (see (4-12)) simultaneously. Thus, the delay of each hop is no more than \( \Theta(\sqrt{N \log N}) \). The total delay is no more than \( \sqrt{2} / \delta r \Theta(\sqrt{N \log N}) = \Theta(N) \).}

E. Cost of routing tables

Define abbreviations and acronyms the first time they are used in the text, even after they have been defined in the abstract. Do not use abbreviations in the title unless they are unavoidable.

**Theorem 4.3.** In Algorithm 3.1, if node \( n_1 \) has at least one neighbor node, then at least \( \left\lceil \frac{|M|}{2} \right\rceil \) directions need not to memorize any routing table entry.

Proof: it is easy to see that in Fig. 4.2, since directions equally divide a plane, the half-plane in the \( n_2 \) side of the dotted line has at least \( \left\lceil \frac{|M|}{2} \right\rceil \) directions (in Fig. 4.2, the dotted line is perpendicular to line \( n_{2}n_{3} \)). We can see that, in Fig. 4.2, the projection of \( n_2 \) on directions is always large than that of \( n_1 \), thus we do not need to store any virtual link between \( n_1 \) and \( n_2 \) for these directions.

\[ \text{Figure 4.2. Illustration of Theorem 4.4} \]

**Theorem 4.4.** In Algorithm 3.1, an upper bound of the number of routing table entries in a node is \( (N-1) \left\lceil \frac{|M|}{2} \right\rceil \).

Proof: there are at most \( \left\lceil \frac{|M|}{2} \right\rceil \) directions needed for a node to record the routing path. Therefore the theorem is proved.

F. Analysis of stability

Link and direction schedule dispatch the packets and decide the working state of links according to the interference constraint. The schedules are required to ensure the network eventually reaches a stable state. If the queue length \( Q_l \) of a link \( l \) is finite, the link \( l \) is called stable. If all links in the network are stable, the network is called stable. Furthermore, the network stability means that the loads are averaged to nodes and the load of each node is in scope of tolerance. Next, we prove the stability of MDR protocol by analysis. Then, in Section 5, we give the experimental results of load balance comparison between MDP and other protocols.

Let \( \lambda = (\lambda_0, \lambda_1, \ldots, \lambda_{n-1}) \) represent the arrival rate vector of the network. The literature [26] defines all arrival rate vectors that a schedule algorithm \( \Pi \) can keep the network stable with as capacity region \( \Lambda_{\text{opt}} \), that is, \( \Lambda_{\text{opt}} = \{ \lambda | \text{schedules} \Pi \text{stable with} \lambda \} \). For a special network, the union capacity region of all schedule algorithms is called optimal capacity region \( \Lambda_{\text{opt}} \). If any schedule algorithm can reach \( \Lambda_{\text{opt}} \), the schedule is called the optimal schedule \( \Pi_{\text{opt}} \). As defined, \( C_l \) is transmitting capacity of link \( l \). Let \( r_l \) denote the workable data rate of link \( l \) in the current state. As easy to see, \( r_l = C_l \) when scheduled and \( r_l = 0 \) when not scheduled. Let \( r = \{r_0, r_1, \ldots, r_{n-1}\} \) denote the workable data rate vector of total \( L \) links in the network. Collection of all workable vectors \( R \) can be written as \( R = \{r | r \text{ workable currently} \} \). According to literatures [26] [33], \( \Lambda_{\text{opt}} \) can be defined as \( \Lambda_{\text{opt}} = \{ \lambda | \lambda \leq r, r \in \text{Conv}(R) \} \), in which \( \text{Conv}(R) \) denotes the convex collection of \( R \). This definition implies that for any arrival data rate vector \( \lambda' \in \Lambda_{\text{opt}} \), there must exist a workable schedule vector \( r' \in \text{Conv}(R), \quad r' = r \).

It is hard to find the optimal solution under the constraint of interference between nodes [34]. To find a sub-optimal solution is a workable way. The capacity region of sub-optimal solution reaches \( \gamma (0 < \gamma < 1) \) times of optimal solutions. That is, if a schedule \( \Pi \) can keep the network stable with any arrival data rate vector \( \lambda \in \gamma \Lambda_{\text{opt}} \), the schedule \( \Pi \) is called a sub-optimal solution. Next, we prove our schedule is a sub-optimal solution.

**Lemma 4.1.** For any arrival data rate vector \( \lambda \in \Lambda_{\text{opt}} \), there must exist a random schedule that make the expect value of schedule vector \( E[r] = \lambda \).
(That is, the schedule can process the data with the arrival rate and do not increase the queue length of each node in the network).

Proof: Under the definition of $\Lambda_{\text{opt}}$, for any $\lambda \in \Lambda_{\text{opt}}$, there must exist a workable schedule vector $r \in \text{Conv}(R)$, $= r$. According to the Carathéodory theorem (The classical theorem in convex geometry), convex set $\text{Conv}(R)$ must have:

$$r = P_1 r_1 + P_2 r_2 + \ldots + P_l r_l \quad (4-13)$$

in (4-13), $r_i \in R (i=0,1,..,L), \sum P_i = 1$. So a random schedule algorithm exists: randomly select a schedule vector $r_i$ from $\{r_1, r_2, \ldots, r_l\}$ with corresponding probability $P_i$. This schedule has $\{r\} = \lambda$. \hfill $\blacksquare$

**Theorem 4.5.** The inequality $\sum_{l=0}^{l \in S_{\text{opt}}} P_l r_i \geq \gamma \sum_{l=0}^{\kappa_{\max}} r_i (\gamma \geq 1/\kappa_{\max})$ is satisfied. In this inequality, $\kappa_{\max} = \max \kappa(l), \kappa(l)$ denotes the maximal count of links that can simultaneously transmit in interference set $I(l)$.

Proof: set the links set of optimal schedule $\pi_{\text{opt}}$ as $S_{\text{opt}}$, and $\pi_{\text{MDR}}$ as $S_{\text{MDR}}$. For $l \in S_{\text{opt}}, l \notin S_{\text{MDR}}$, there exists a link set $E(l) \subseteq S_{\text{MDR}}$ and $E(l) \subseteq I(l)$.

In $E(l)$, there exists link $l$ and has $Q_i r_i \leq Q_i r_i$. (Recalling $Q_i$ is the weight for link schedule of $\text{MDR}$ and $Q_i r_i$ is the weight for direction schedule of $\text{MDR}$ to smooth the flow). So for set $S_{\text{opt}}$, there must exist set $S_{\text{opt}}' = \{l_i \in S_{\text{MDR}} | i = 0,1,..,S_{\text{MDR}}| -1\}$.

In $S_{\text{MDR}}$, some links are the same. Randomly select a group of same links $l_{i_0}^i = l_i = \ldots = l_{i_k}^i$, totally $k$. Then, there will be $k$ different links $\{l_0^i, l_1^i, .., l_k^i \} \in S_{\text{MDR}}$ and $l_i \in I(l_i)$ for $i=0,1,..,k$. Obviously, $k \leq \kappa_{\max}$ as defined. Otherwise, the case that any two links $l_i, l_j$ in $S_{\text{MDR}}$ are the same will contradict to the optimal schedule itself. Another link $l_{i_q}^i$ with $Q_i r_i \leq Q_i r_i$ can be found in this case.

As for $\forall l \in S_{\text{MDR}}, l \notin S_{\text{opt}}$, there exists $Q_i r_i \leq Q_i r_i, l \in S_{\text{MDR}}$. We can get the inequality:

$$\sum_{i=0}^{i \in S_{\text{opt}}} Q_i r_i \leq \sum_{i=0}^{i \in S_{\text{MDR}}} Q_i r_i \quad (4-14)$$

Let $S_{\text{MDR}}^\prime$ denote the set that selects distinct links from $S_{\text{MDR}}$. Obviously, the following inequality can be established:

$$\sum_{i=0}^{i \in S_{\text{MDR}}^\prime} Q_i r_i \leq \gamma \sum_{i=0}^{i \in S_{\text{MDR}}^\prime} Q_i r_i \quad (4-15)$$

According to the definition, for any $l \in S_{\text{MDR}}^\prime$, there is $l \notin S_{\text{MDR}}$. That is, $S_{\text{MDR}}^\prime \subseteq S_{\text{MDR}}$. Then, we have:

$$\sum_{i=0}^{i \in S_{\text{MDR}}^\prime} Q_i r_i \leq \gamma \sum_{i=0}^{i \in S_{\text{MDR}}^\prime} Q_i r_i \quad (4-16)$$

Combine (4-14), (4-15) and (4-16), we get the result:

$$\sum_{i=0}^{i \in S_{\text{MDR}}^\prime} Q_i r_i \geq \gamma \sum_{i=0}^{i \in S_{\text{MDR}}^\prime} Q_i r_i \quad (4-17)$$

In (4-17), $\gamma \geq 1/\kappa_{\max}$. \hfill $\blacksquare$

**Theorem 4.6.** $\Pi_{\text{MDR}}$ is a sub-optimal schedule. For any arrival data rate, if $+ (\ v = (v_0, v_1, \ldots, v_L), v_0 > 0, i = 0,1,..,L - 1 \ )$, Then, $\Pi_{\text{MDR}}$ can guarantee the network stable and has the supremum limit $\lim_{s \to \infty} \sum_{i=0}^{b_i} E_i(Q_i(z)) < B/\varepsilon$ ($B$ is constant vector).

Proof: in state $s$, any link $l$ has $Q_i(s+1) = \max \{Q_i(s) - r_i(s), 0\} + \lambda_i(s)$. As defined, $r_i(s)$ is the schedule capacity and $\lambda_i(s)$ is the arrival data rate of link $l$ in state $s$. As easy to verify, there is $Q_i(s+1) - Q_i(s) \leq r_i(s) + \lambda_i(s) - 2Q_i(s)(\tau_i(s) - \lambda_i(s))$.

Define Lyapunov function [35]: $L(Q(s)) = \sum_{i=0}^{i \in S_{\text{MDR}}}$. We have the following inequality:

$$L(Q(s+1) - L(Q(s)) \leq \sum_{i=0}^{i \in S_{\text{MDR}}} (r_i^2(s) + \lambda_i^2(s))$$

(4-18)

Assume $r_{\max} = \max r_i(s)$ and $\lambda_{\max} = \max \lambda_i(s)$ for $l = 0,1,..,L, s = 0,1,..,s$. Then, we have

$$\sum_{i=0}^{i \in S_{\text{MDR}}} (r_i^2(s) + \lambda_i^2(s)) \leq L(r_i(s) + \lambda_i(s)) = B \quad (4-19)$$

Define Lyapunov drift function [35] as:

$$\Delta Q(s) = E_{\text{MDR}}(L(Q(s+1)) - L(Q(s))) \quad (4-20)$$

Combine (4-18)-(4-19) and (4-20), we have

$$\Delta Q(s) \leq B + 2\sum_{i=0}^{i \in S_{\text{MDR}}} (Q_i(s) r_i(s))$$

(4-21)

Refer to Lemma 4.1, we know if $\lambda(s) + \varepsilon \in \gamma L$, there
exists an optimal random schedule \( \Pi_{opt} \) that have:

\[
E_{opt}(r(s)) = \left( \lambda(s) + \varepsilon \right) / \gamma
\]  

(4-22)

Also, \( E(r(s)) = \sum_{i=0}^{L} E(r_i(s)) \). According to Theorem 4.5, \( \sum_{i=0}^{L} Q_i r_i \geq \gamma \sum_{i=0}^{L} E(r_i(s)) (\gamma \geq 1 / k_{max}) \), we have

\[
E_{MDR}(Q(s)r(s)) \geq \gamma E_{opt}(Q(s)r(s)) \geq \sum_{i=0}^{L-1} Q_i (\lambda(s) + \varepsilon) \]  

(4-23)

Substitute (4-23) in (4-21), we get:

\[
\Delta Q(s) = B - 2\varepsilon \sum_{i=0}^{L-1} Q_i(s)
\]  

(4-24)

Refer to the Lyapunov stability theorem [35], we get the result:

\[
\lim_{s \to +\infty} \sup_{s} 1/ s \sum_{\varepsilon=0}^{s} E(Q(\varepsilon)) < B/\varepsilon
\]  

(4-25)

Therefore, Theorem 4.6 is proved. ■

V. EVALUATIONS

In this section, we will evaluate and compare our algorithm (Class II) with Shortest-Path Routing (SPR) algorithm (Class I), Hyperbolic Geographic Routing (HGR) algorithm (Class III) [13], and Randomized Local Load Balancing (RandLLB) algorithm (Class IV) [25]. These algorithms are classical methods belonging to different categories in Table 2.1. We perform the experiments on Matlab.

A. Performance Metrics

1) Throughput

Throughput is evaluated by packet delivery ratio \( R_T \). First, the transmission success rate \( T_\theta \) (where \( \theta \) can be SPR, HGR and RandLLB) is defined as \( T_\theta = (n_a/n_i) \times 100\% \), in which \( n_a \) is the number of packets that arrive at destinations; \( n_i \) is the total number of packets sent out. Then we have

\[
R_T = \frac{T_{MDR}}{T_\theta} \times 100\%
\]  

(5-1)

2) End-to-end Delay

We use delay ratio \( R_D \) to present end-to-end delay (This includes possible delays such as queuing delays at the interface, propagation delays and transmission delays). After all packets reached their destinations, we record the delay of each packet and denote the delay of the packet \( i \) as \( d_i \). Assuming there are totally \( P \) packets generated, the average delay \( D_\theta \) for all packets is defined as \( D_\theta = \frac{1}{P} \sum_i d_i \). Then we have

\[
R_D = \frac{D_{MDR}}{D_\theta} \times 100\%
\]  

(5-2)

3) Load Balance

We use load ratio \( R_L \) to measure the property of load balance. In a given time period, we record the total number of packets processed by node \( i \) as \( l_i \). The average number of packets processed by each node is defined as \( E_\theta = \frac{1}{P} \sum_i l_i \). Then, the variance of the number of packets processed by each node is defined as \( V_\theta = \frac{1}{P} \sum_i (l_i - E_\theta)^2 \) which indicates how smoothly packets will distribute on all nodes. When \( V_\theta \) is smaller, it means that packet flows distribute more smoothly in the network. Then we have

\[
R_L = \frac{V_{MDR}}{V_\theta} \times 100\%
\]  

(5-3)

4) Routing Table Size

We use space ratio \( R_S \) to evaluate storage cost (i.e. the routing table size). Denoting the routing table size of different algorithms as \( S_\theta \), then we have

\[
R_S = \frac{S_{MDR}}{S_\theta} \times 100\%
\]  

(5-4)

B. Simulation Environment

We use following settings in our simulations: in a given region, there are 100 nodes randomly deployed. For each trial of experiments, 10 source-destination pairs of packet flows among nodes are randomly generated. For each flow, the number of generated packets varies from 60 to 300 (in 1000 time slots). For the capability of network nodes, we assume that each node can process 30 packets per slot. For ease of evaluating, we also assume that all nodes have a queue with infinite length and there is no packet loss due to queuing. The simulation will terminate when all packets have been delivered to their destinations. For the direction assignment, we use the even assignment in Fig. 3.1.

C. Comparison of Throughput

The comparison of throughput is showed in Fig. 5.1 (a), (b), and (c). We can see that: 1) when the number of packets increases, our algorithm performs better than other algorithms if more directions are available (here 3 IDPs are enough); 2) The performance of our...
algorithm is directly proportional to the number of directions available. That is, the more directions used, the better our algorithm is. Especially, when 7 or more IDPs are selected, our algorithm performs at least 20% better than the SPR algorithm, 30% better than the HGR algorithm. For the reason of packets replication and destination unreachable, delivery ratio of RandLLB is very high.

E. Comparison of Load Balance

The comparison of load balance is showed in Fig. 5.3. From (5-4), we know that if $R_L$ is less than 100%, it means our algorithm is better than other algorithms in respect of load balance. Fig. 5.3 shows that our algorithm has a smaller variance value, which means that packet flows distribute more smoothly than that of the SPR, HGR and RandLLB algorithm when enough directions are used. In addition, we can see that when more directions are used, the load is more balanced in our algorithm.

D. Comparison of End-to-End Delay

The comparison of end-to-end delay is showed in Fig. 5.2 (a), (b), and (c). We can see that, first, when the number of packets increases, our algorithm performs better (i.e. end-to-end delay is smaller) if enough directions are available (here 3 IDPs are enough); second, the end-to-end delay of our algorithm is about 65% of the HGR algorithm, 70% of that of the SPR and RandLLB algorithm.
Comparison of Routing Table Size

Figure 5.4 gives the comparison of routing table sizes for all algorithms. For our algorithm, when available directions increase, the size of the routing table also increases. When 8 IDPs are used, the routing table size of our algorithm is approximately twice of that of the SPR algorithm. When 4 IDPs are used, the routing table size of our algorithm is almost equal to that of the SPR algorithm. Since HGR and RandLLB algorithms are both based on greedy routing without routing tables, we only compare with the SPR algorithm here. In exchange, these two algorithms pay the price for the worse performance on delivery ratio, delay ratio and load balance.

Remarks. As showed in the results, the tradeoff of our algorithm is to use more space to gain better performance on throughput, end-to-end delay and load balance. Our approach can obtain the remarkable improvement with acceptable space cost. For example, when we use seven IDPs, there is at least 20% improvement on throughput, end-to-end delay and load balance with the double of space cost. Besides, devices are equipped with more storage with the rapid growth of storage technologies. Therefore, our algorithm is feasible in practice.

VI. CONCLUSIONS

In this paper, we have discussed a new routing protocol, MDR. Compared to current geographical routing protocols, MDR uses direction as a metric for routing instead of using location as metric.

With the newly proposed direction metric, MDR provides a new solution on the issue of avoiding packets lost in local minimum that a geographical routing protocol should concern.

At the same time, MDR analyzes its performance on the issues of throughput, end-to-end delay, load balance and network stability that previous geographical routing protocol often ignored. As proved, every source-destination pair can achieve a data rate $\Theta\left(\frac{1}{\sqrt{\text{NlogN}}}\right)$ simultaneously, while end-to-end delay is directly proportional to the network size. In simulation, we show that our method achieves good performance with acceptable space cost.

When the node density is sparse, the paths in different directions may have overlap parts. In the future, we will analyze the affect of the overlap parts of the paths to find the routing bottlenecks; then, to improve the MDR algorithm by discovering the alternative paths and deploying extra nodes along the bottleneck paths. Besides, we can further reduce the size of the routing table by applying the overlap analysis.

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