Current Situation:

Every permutation could be routed on a permutation network and meshes. The number of steps is proportional to the diameter. The algorithm was centralized and needed global knowledge about the sources and destinations of all packets.

We now want to devise local-control algorithms. Each node decides on the next step by some local information.
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The number of steps $T$ taken by an algorithm to deliver all packets is referred to as routing time.
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- **Example**: bit-fixing paths on the hypercube
Lower Bound by Borodin and Hopcroft

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The time complexity for permutation routing under this paradigm is lower bounded by $\Omega(\sqrt{n/\Delta})$, which is polynomial in $n$.

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Proof of the lower bound by Borodin and Hopcroft
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- For $v \in V$, let $\mathcal{W}_v = \{P_{v,u} \mid u \in V\}$.
- For a positive number $t$, a node $v \in V$, and an edge $e \in E$, we say that $e$ is $t$-popular for $v$ if at least $t$ paths from $\mathcal{W}_v$ contain $e$. 
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Outline of the proof:

- First, we prove a lemma showing that, for any given node $v \in V$, there are many edges that are "quite popular" for $v$. 
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- Given this, we will be able to construct a permutation $\pi$ such that $t$ of the paths selected by $\pi$ contain $e^*$, which proves the lower bound.
Proof of the lower bound by Borodin and Hopcroft

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One Lemma for the Proof of the lower bound

Lemma

\( \forall v \in V \text{ and } t \leq (n - 1)/\Delta : A_v(t) \geq \frac{n}{2\Delta t} \).

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**Lemma**

∀v ∈ V and t ≤ (n − 1)/Δ : A_v(t) ≥ \( \frac{n}{2\Delta t} \).

**Proof of lemma:**

- Let Q ⊆ V be the set of nodes from which there is a path to v that contains only edges that are t-popular for v.
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- Let \( Q \subseteq V \) be the set of nodes from which there is a path to \( v \) that contains only edges that are \( t \)-popular for \( v \).

- Let \( L = V - Q \) and \( B = E \cap (L \times Q) \), that is, \( B \) is the set of those edges connecting a node in \( L \) with a node in \( Q \).
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|B| \cdot (t − 1) ≥ |L| because, for each node u ∈ L, the path P_v,u leads through at least one edge in B and these edges are not t-popular so that each of them can be contained in at most t − 1 paths from W_v.
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**Lemma**

\[\forall v \in V \text{ and } t \leq \frac{(n - 1)}{\Delta} : A_v(t) \geq \frac{n}{2\Delta t}.\]

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- It holds
  - \(|B| \cdot (t - 1) \geq |L|\) because, for each node \( u \in L \), the path \( P_{v,u} \) leads through at least one edge in \( B \) and these edges are not \( t \)-popular so that each of them can be contained in at most \( t - 1 \) paths from \( W_v \).
  - \(|B| \leq \Delta |Q| \) as each node in \( Q \) has at most \( \Delta \) incident edges.
Proof of the lemma

- Combining the two equations, we obtain

\[ \Delta |Q| (t - 1) \geq |L| = n - |Q|, \]
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Next we will show \(|Q| \leq 2A_v(t)| which completes the proof of the lemma as it implies

\[ A_v(t) \geq \frac{|Q|}{2} \geq \frac{n}{2\Delta t}. \]
Proof of the lemma

Let $E'$ denote the set of edges that are $t$-popular for $v$. To complete the proof of the lemma, we have to show $|Q| \leq 2|E'| = 2A_v(t)$. 
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such that at least one of the edges incident to $v$ is contained in at least $(n - 1)/\Delta \geq z$ paths from $\mathcal{W}_v$. 

Show: $|Q| \leq 2A_v(t)$
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Therefore, there is at least one edge that is $t$-popular for $v$. 

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Given that $E'$ is non-empty, each node in $Q$ is incident to an edge in $E'$.

Consequently, $|Q| \leq 2|E'|$ as each of the edges in $E'$ is incident to at most two nodes from $Q$. 
Proof of the lower bound by Borodin and Hopcroft

Show: \( \exists e^*: \) \( e^* \) is \( t \)-popular for \( t \) different nodes, for \( t = \Omega(\sqrt{n}/\Delta) \).

- Our next goal is to show that there exists an edge \( e^* \) that is \( t \)-popular for \( t \) nodes where \( t = \Omega(\sqrt{n}/\Delta) \).
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$$\sum_{e \in E} A_e(t) = \sum_{e \in E} \sum_{v \in V} A_{e,v}(t) = \sum_{v \in V} \sum_{e \in E} A_{e,v}(t) = \sum_{v \in V} A_v(t) \geq \frac{n^2}{2\Delta t},$$

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where the inequality follows from the lemma.
- Because of the “pigeonhole principle”, there has to exist an edge \( e^* \in E \) such that

\[
A_{e^*}(t) \geq \left\lceil \frac{n^2}{|E| \cdot 2\Delta t} \right\rceil \geq \left\lceil \frac{n}{2\Delta^2 t} \right\rceil,
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where the last step follows from \( |E| \leq \Delta n \).
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- Observe that \( t = \sqrt{n}/(\sqrt{2}\Delta) \) implies \( t \leq (n - 1)/\Delta \), for any \( n \geq 2 \),
- so that the assumption about \( t \) that we made in the lemma is satisfied.
Proof of the lower bound by Borodin and Hopcroft

Show: \( \exists e^*: e^* \text{ is } t\text{-popular for } t \text{ different nodes, for } t = \Omega(\sqrt{n}/\Delta). \) We have \( A_{e^*}(t) \geq \lceil \frac{n}{2\Delta^2 t} \rceil \).

Next we choose \( t \) such that \( t = \frac{n}{2\Delta^2 t} \), that is, we set \( t = \sqrt{n}/(\sqrt{2}\Delta) \).

- Observe that \( t = \sqrt{n}/(\sqrt{2}\Delta) \) implies \( t \leq (n - 1)/\Delta \), for any \( n \geq 2 \),
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- For this choice of \( t \), our analysis gives
Proof of the lower bound by Borodin and Hopcroft

Show: ∃e*: e* is t-popular for t different nodes, for t = Ω(√n/∆). We have A_{e*}(t) ≥ ⌈n/2Δ^2⌉.

- Next we choose t such that t = \frac{n}{2Δ^2}, that is, we set t = \frac{\sqrt{n}}{(\sqrt{2})}. 
- Observe that t = \frac{\sqrt{n}}{(\sqrt{2})} implies t \leq (n - 1)/Δ, for any n ≥ 2, 
- so that the assumption about t that we made in the lemma is satisfied.

For this choice of t, our analysis gives

- A_{e*}(t) ≥ \left\lceil \frac{n}{2Δ^2 \sqrt{n}/(\sqrt{2Δ})} \right\rceil = \left\lceil \frac{n\sqrt{2Δ}}{2Δ^2 \sqrt{n}} \right\rceil = \left\lceil \frac{\sqrt{n}}{\sqrt{2Δ}} \right\rceil.
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  - \( e^* \) is \( \lceil t \rceil \)-popular for \( \lceil t \rceil \) nodes, where \( t = \sqrt{n}/(\sqrt{2}\Delta) \).
Proof of the lower bound by Borodin and Hopcroft

Construct a permutation $\pi$ such that $t$ of the paths selected by $\pi$ contain $e^*$. Finally, we construct a permutation $\pi$ such that $\lceil t \rceil$ of the paths selected by $\pi$ contain $e^*$:

- Let $V'$ denote a set of $\lceil t \rceil$ nodes for which $e^*$ is $\lceil t \rceil$-popular.
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- W.l.o.g., \( V' = \{1, \ldots, \lceil t \rceil \} \).
- For every \( v \in V' \), there exists a subset \( U_v \subseteq V \) of cardinality \( \lceil t \rceil \) such that, for every \( u \in U_v \), the path \( P_{v,u} \) contains \( e^* \).
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- For $v = 1$ to $\lceil t \rceil$, set $\pi(v) = u$ where $u$ is chosen arbitrarily from $U_v \setminus \{\pi(1), \ldots, \pi(v-1)\}$.
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By our construction, $\pi$ and $e^*$ satisfy the properties described in the theorem.
Application to the hypercube and Goal

- For the $d$-dimensional hypercube with $n = 2^d$ nodes, the lower bound of Borodin and Hopcroft implies a lower bound of $\Omega(\sqrt{n} / \log n)$ for permutation routing.
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- Our goal is to devise a distributed permutation routing algorithm with time complexity $O(\log n)$. 
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- Consequently, when using bit-fixing paths the time complexity for permutation routing is $\Omega(\sqrt{n})$.

- Our goal is to devise a distributed permutation routing algorithm with time complexity $O(\log n)$.

- This will take some time.
Outline of the approach

- We build a dynamic system of storage devices supporting the addition and removal of storage devices using dynamic hashing:

\[ \text{1} \text{ independently, uniformly at random} \]
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- We build a dynamic system of storage devices supporting the addition and removal of storage devices using dynamic hashing:
  - devices are mapped i.u.r.\(^1\) to the ring \([0, 1)\), that is, each device \(i\) gets assigned a random address \(a(i) \in [0, 1)\)

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- We assume an idealistic hash function, that is, the hash values are real numbers chosen i.u.r. from \([0, 1)\).

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Definition of successors

- Let $V$ be the set of storage devices at some point of time, and let $n = |V|$. 
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- For address \( A \in [0, 1) \), define

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\text{succ}(A) = \begin{cases} 
\arg\min \{ a(i) \geq A \mid i \in V \} & \text{if } \exists i \in V : a(i) \in [A, 1), \\
\arg\min \{ a(i) \geq 0 \mid i \in V \} & \text{otherwise.}
\end{cases}
\]

\[
\text{pred}(A) = \begin{cases} 
\arg\max \{ a(i) < A \mid i \in V \} & \text{if } \exists i \in V : a(i) \in [0, A), \\
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- Object $x \in U$ is mapped to device $\text{succ}(h(x))$. 
Quality of the load balancing

- The quality of the load balancing depends on the distribution of the sizes of the ring for which the storage devices are responsible.
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**Definition (weight of a device)**

For device $i \in V$, define the weight of device $i$ by

$$W_i = \begin{cases} 
  a(i) - a(pred(a(i))) & \text{if } a(pred(a(i))) < a(i), \\
  1 - (a(pred(a(i))) - a(i)) & \text{otherwise.}
\end{cases}$$

Let $W = \max_{i \in [n]} W_i$. 
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- Ideally, we would have $W = W_0 = W_1 = \ldots = W_{n-1} = \frac{1}{n}$. 
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Let \( W = \max_{i \in [n]} W_i \).

- Ideally, we would have \( W = W_0 = W_1 = \ldots = W_{n-1} = \frac{1}{n} \).
- We will show that \( W = O\left(\frac{\log n}{n}\right) \), w.h.p.\(^2\)

\(^2\)The term "w.h.p." abbreviates "with high probability and means with probability at least \( 1 - n^{-\alpha} \), for any constant \( \alpha > 0 \).
Lemma

Let $T \subseteq [0, 1)$ and $t = |T|$ the mass (length) of $T$. Suppose that $M$ points are chosen i.u.r. from $[0, 1)$. The probability that none of these points is from $T$ is at most $e^{-tM}$.

Proof:

$$\Pr[\text{no point in } T] = (1 - t)^M = ((1 - t)^{1/t})^tM \leq e^{-tM}$$

as, for every $x > 0$, it holds $(1 - \frac{1}{x})^x \leq \frac{1}{e}$. 
Quality of the load balancing

Theorem

\[ W = O\left(\frac{\log n}{n}\right), \text{ w.h.p.} \]

Proof:

- Fix \( j \in V \). Suppose \( j \)'s address \( a(j) \) is fixed arbitrarily.

Pr [no point in \( T \) ] \( \leq e^{-tM} \).
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- Fix \( j \in V \). Suppose \( j \)'s address \( a(j) \) is fixed arbitrarily.
- A necessary condition for the event \( W_j \geq t \), \( t \in [0,1) \), is that no addresses of the other \( n - 1 \) devices falls into the interval from \( a(j) - t \) to \( a(j) \).
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- For any \( \alpha > 0 \),

\[
\Pr \left[ W_j \geq 2(\alpha + 1) \frac{\ln n}{n} \right] \leq e^{-2(\alpha + 1) \frac{\ln n}{n} (n-1)}
\]

\[
\leq e^{-2(\alpha + 1) \ln n} = n^{-(\alpha + 1)}
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- for any \( \alpha > 0 \),
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  \leq e^{-(\alpha+1)\ln n} = n^{-(\alpha+1)}
  \]
- and, hence,
  \[
  \Pr \left[ W \geq 2(\alpha + 1)\frac{\ln n}{n} \right] \leq \sum_{j \in V} \Pr \left[ W_j \geq 2(\alpha + 1)\frac{\ln n}{n} \right] \leq n^{-\alpha}.
  \]
Improved quality of the load balancing

- In order to improve the load balancing, we use $k$ virtual nodes for each device. Let $V'$ denote the set of $kn$ "virtual nodes. We have $W = O\left(\frac{\log n}{n}\right)$, w.h.p.
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- For address \( A \in [0,1) \), re-define

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succ(A) = \begin{cases} 
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- Object \( x \in U \) is mapped to node \( succ(h(x)) \) and stored on the device to which this node belongs.
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- Let $W_i$ denote the weight of device $i$, i.e., the sum of the lengths of the intervals corresponding to $i$'s nodes, and $W = \max_{i \in [n]} W_i$. 

Theorem

For any $k \geq 1$, $W = \frac{1}{n} \cdot O\left(1 + \frac{\log n}{k}\right)$, w.h.p.
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**Theorem**
For any $k \geq 1$, $W = \frac{1}{n} \cdot O(1 + \frac{\log n}{k})$, w.h.p.

**Corollary**
If $k \geq \log n$ then $W = O(\frac{1}{n})$, w.h.p.
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**Proof of the Theorem:**

- Consider device \( j \) and suppose the address of the \( k \) nodes of this device are fixed arbitrarily.
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**Proof of the Theorem:**

- Consider device \( j \) and suppose the address of the \( k \) nodes of this device are fixed arbitrarily.
- For any \( t \in [0, 1) \), we want to upper-bound \( \Pr[W_j \geq t] \).
Improved quality of the load balancing

**Exact condition:**

The event $W_j \geq t$ happens if and only if there are $k$ intervals left of the $k$ addresses of $j$’s nodes so that

- these intervals have a total length of $t$, and
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Exact condition:

The event $W_j \geq t$ happens if and only if there are $k$ intervals left of the $k$ addresses of $j$’s nodes so that

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In order to be able to enumerate all possibilities for choosing these $k$ intervals, we look at a slightly stronger necessary condition for the event $W_j \geq t$. 

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**Necessary condition:**

If the event $W_j \geq t$ happens then there are $k$ intervals left of the $k$ addresses of $j$’s nodes so that

- the length of each of these intervals is a multiple of $\frac{1}{kn}$.
Improved quality of the load balancing

**Exact condition:**

The event $W_j \geq t$ happens if and only if there are $k$ intervals left of the $k$ addresses of $j$’s nodes so that

- these intervals have a total length of $t$, and
- none of the other $k(n - 1)$ nodes have an address that falls into these intervals.

**Necessary condition:**

If the event $W_j \geq t$ happens then there are $k$ intervals left of the $k$ addresses of $j$’s nodes so that

- the length of each of these intervals is a multiple of $\frac{1}{kn}$
- these intervals have a total length of $t'$ where $t'$ is the largest multiple of $\frac{1}{kn}$ such that $t' \leq t - \frac{1}{n}$, and
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- The number of possibilities to choose these intervals corresponds to the number of possibilities to choose $k$ integers $q_1, \ldots, q_k$ such that $\sum_{i=1}^{k} q_i = q$, for $q = t'kn$. 

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- Thus, the number of possibilities to choose the $q_i$’s and, hence, the intervals is at most

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- Now $q + k = t'kn + k \leq (t - \frac{1}{n})kn + k = tkn$, so that this number is at most
  \[
  \left( \frac{etkn}{k} \right)^k = (etn)^k.
  \]
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- Once the intervals are fixed, the probability that these intervals with a total length of \( t' \geq t - \frac{2}{n} \) are not hit by one other \( k(n - 1) \) addresses is at most

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- Now choose $t = \beta \frac{n}{n}$, where the value for $\beta$ will be specified later.

- This gives $etn = e\beta$ and
  $$\left(t - \frac{2}{n}\right)(n - 1) \geq \left(t - \frac{2}{n}\right) \frac{n}{2} = \frac{\beta}{2} - 1$$

  assuming $n \geq 2$. 
Improved quality of the load balancing

Consequently,

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\Pr \left[ W_i \geq \frac{\beta}{n} \right] \leq \left( e^\beta \cdot e^{-\beta/2+1} \right)^k
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= \left( e^\beta \cdot e^{-\beta/2+1} \cdot \left( \frac{4}{3} \right)^\beta \right)^k \left( \frac{3}{4} \right)^{\beta k}.
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Now observe that \( e^\beta \cdot e^{-\beta/2+1} \cdot \left( \frac{4}{3} \right)^\beta \) decreases exponentially in \( \beta \) since \( e^{\frac{1}{2}} \geq \frac{4}{3} \). For \( \beta \geq 25 \), this term is less than 1. Consequently,

\[ \Pr \left[ W_i \geq \frac{\beta}{n} \right] \leq \left( \frac{3}{4} \right)^{\beta k} \leq \left( \frac{3}{4} \right)^{(\alpha+1) \log_4 \frac{4}{3} n} = n^{-(\alpha+1)}, \]

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It follows $\Pr \left[ W \geq \frac{\beta}{n} \right] \leq n^{-\alpha}$, for $\beta = O(1 + \frac{\log n}{k})$, which proves the theorem.
Now we connect the nodes from the consistent hashing scheme by an overlay network called Chord running on top of the Internet.
Overlay network

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- We say that node $v$ has a link to node $u$ if $u$’s IP address is stored in the finger table of $v$. ³

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- The Chord network allows that devices enter and leave the system dynamically and supports the efficient search for data objects.

---

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Definition of the Chord edges

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- Observe that the set of links is finite. For $v \in V$, let $d(v)$ denote the smallest integer such that
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  \[ \forall i \in \mathbb{N}, i \geq d(v) : e(v, i) = (v, \text{succ}(a(v))) \]
- The outdegree of $v$ is at most $d(v)$. Let $D = \max\{d(v) \mid v \in V\}$. 
Upper-bounding the outdegree

Theorem

\[ D = O(\log n), \text{ w.h.p., where } n = |V|. \]

Proof:
- Consider any node \( v \in V \).
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- All edges \( e(v, i) \) with \( i \geq d(v) \) point to \( \text{succ}(a(v)) \).
- In particular, it holds \( 2^{-d(v)} \leq \ell(v) \), which gives

\[
d(v) = \left\lceil \log \left( \frac{1}{\ell(v)} \right) \right\rceil.
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- if at least one of the other $n - 1$ nodes falls into the interval $[a(v), a(v) + \beta)$ which, for each of these nodes, happens with probability $\beta$.  

Upper-bounding the outdegree

Now let $\alpha > 0$ be chosen arbitrarily. We obtain

$$
\Pr \left[ d(v) \geq (\alpha + 3) \log n \right] \leq \Pr \left[ \log \left( \frac{1}{\ell(v)} \right) \geq (\alpha + 3) \log n \right]
$$

$$
\leq \Pr \left[ \log \left( \frac{1}{\ell(v)} \right) > (\alpha + 2) \log n \right]
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Hence, the probability that there exists a node $v \in V$ for which $d(v) \geq (\alpha + 3) \log n$ is at most $n^{-\alpha}$. \qed
Routing in Chord

- Suppose a node \( v \) (or the device corresponding to \( v \)) wants to access a data object \( x \).
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The object can be found by applying the following routing algorithm:

- First, $v$ checks whether $\text{succ}(h(x)) = v$. If yes, then stop.
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- The number of hops needed for finding an object is at most \( D \) and, thus, \( O(\log n) \), w.h.p., because the index of the outgoing links is increasing with every hop on the routing path.
Oblivious Randomized Routing

Definition

- One specifies a path system $\mathcal{W}$ containing a set of paths $W_{u,v}$ from $u$ to $v$
Oblivious Randomized Routing

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- One specifies a path system $\mathcal{W}$ containing a set of paths $W_{u,v}$ from $u$ to $v$
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One specifies a path system $\mathcal{W}$ containing a set of paths $W_{u,v}$ from $u$ to $v$ together with a probability distribution $D_{u,v} : W_{u,v} \rightarrow [0, 1]$, for every possible source-destination pair $(u, v) \in V^2$. 
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- For any two nodes $u, v \in V, u \neq v$, one specifies two alternative paths, that is, $|W_{u,v}| = 2$. 

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Example:

- For any two nodes $u, v \in V, u \neq v$, one specifies two alternative paths, that is, $|W_{u,v}| = 2$.
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Example:

- For any two nodes $u, v \in V, u \neq v$, one specifies two alternative paths, that is, $|W_{u,v}| = 2$.
- Let $D_{u,v}$ denote the uniform distribution on $W_{u,v}$.
- When sending a packet from $u$ to $v$ choose $P \in W_{u,v}$ with probability $D_{u,v}(P) = 1/2$. 
Packet Scheduling Problem and Scheduling Policies

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- **Input**: collection of paths $\mathcal{P}$, one for each packet.
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- FTG (Farthest-to-go)
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A scheduling policy is called **greedy** if a packet $p$ has to wait in a step $t$ before using the next edge $e$ on its path only because there is another packet $p'$ using $e$ in this step.
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- FTG (Farthest-to-go)
- Random Rank (as defined later)

- A scheduling policy is called **greedy** if a packet \( p \) has to wait in a step \( t \) before using the next edge \( e \) on its path only because there is another packet \( p' \) using \( e \) in this step.

- We say that \( p \) is delayed by \( p' \) at edge \( e \) in time step \( t \).
Congestion and Dilation

Definition (Dilation)

The dilation $D$ of a path collection $\mathcal{P}$ is the length (number of edges) on the longest path in $\mathcal{P}$.

Definition (Congestion)

The congestion $C$ of a path collection $\mathcal{P}$ is the maximum number of paths from $\mathcal{P}$ that share the same edge (in the same direction).
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- For a (directed) edge \( e \in E \), \( C(e) \) denotes the number of paths from \( \mathcal{P} \) that contain \( e \).
- The congestion is thus defined by \( C = \max_{e \in E} C(e) \).
Trivial bounds on the routing time

Observation (Lower Bound)

*The routing time needed by any scheduling policy is at least\* \(\max\{C, D\} = \Omega(C + D)\) because*

- there is a packet which has path length \(D\) and thus needs at least \(D\) steps to reach its destination, and
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Observation (Upper Bound)

The routing time needed by any greedy scheduling policy is at most \( C \cdot D \) steps because each packet can be delayed at most for \( C - 1 \) steps on each edge on its routing path.
Valiant’s trick

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The node $v_p$ is thus used as intermediate destination.
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Observe that Valiant’s trick follows the paradigm of randomized oblivious routing.
In the following, we present an analysis of phase 1.
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**Lemma**

*The congestion $C$ in phase 1 (phase 2) is $O(\log n / \log \log n)$, w.h.p.*
Proof of the lemma

- Let $e$ be an edge of dimension $i$, i.e., an edge that flips the $i$-th bit.
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- Fix any node in $IN(e)$. The path of the packet starting at $v$ contains $e$ if the packet’s intermediate destination is in $OUT(e)$.
- As intermediate destinations are picked uniformly at random

$$\Pr[v's\ packet\ traverses\ e] = \frac{|OUT(e)|}{n} = \frac{2^i}{2^d} = 2^{i-d}.$$
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- Let $C(e)$ be a random variable describing the congestion at edge $e$, i.e., $C(e)$ is the number of paths containing $e$.
- Let $k$ be any natural number.

\[
\Pr[\exists X \subseteq IN(e), |X| = k : A(X, e)] \\
\leq \sum_{X \subseteq IN(e), |X| = k} \Pr[A(X, e)] \\
= \sum_{X \subseteq IN(e), |X| = k} \left(2^{i-d}\right)^k \\
= \binom{|IN(e)|}{k} \left(2^{i-d}\right)^k.
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Binomial coefficients can be estimated by

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\left( \frac{a}{b} \right)^b \leq \binom{a}{b} \leq \left( \frac{e \cdot a}{b} \right)^b,
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where \( e = 2.71 \ldots \) is the Eulerian number.
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- This gives

\[
\Pr [C(e) \geq k] \leq \left( \frac{e|\ln(e)|}{k} \right)^k \left( 2^{i-d} \right)^k
\]

\[
= \left( \frac{e^{d-i-1}}{k} \right)^k \left( 2^{i-d} \right)^k
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= \left( \frac{e}{2k} \right)^k.
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The congestion is defined to be $C = \max\{C(e) | e \in E\}$.

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\Pr [C \geq k] = \Pr [\exists e \in E : C(e) \geq k] \\
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The last bound follows from $|E| \leq dn \leq n^2$ and $\frac{e}{2k} \leq \frac{1}{2}$, where we assume $k \geq 3$. 
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Now we choose \(k\) such that \(\Pr[C \geq k] \leq n^{-\alpha}\), for constant \(\alpha > 0\).

In particular, we set \(k = \lceil (\alpha + 2) \log n \rceil \geq 3\) which gives

\[
\Pr[C \geq k] \leq n^2 2^{-(\alpha+2)\log n} \leq n^2 n^{-(\alpha+2)} = n^{-\alpha},
\]

which shows \(C = O(\log n)\), w.h.p.
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$$k = \max \left\{ \frac{e}{2} \sqrt{d}, 2(\alpha + 2) \frac{d}{\log d} \right\} = O \left( \frac{\log n}{\log \log n} \right)$$
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  $$\Pr [C \geq k] \leq n^2 \left( \frac{e}{2k} \right)^k \leq n^2 \left( \frac{1}{\sqrt{d}} \right)^k \leq n^2 \left( \left( \frac{1}{\sqrt{d}} \right)^{\frac{2}{\log d}} \right)^{(\alpha+2)d}$$
  
  $$= n^2 \left( \frac{1}{2} \right)^{(\alpha+2)d} = n^2 \cdot n^{-\alpha} = n^{-\alpha}.$$
Congestion of $h$-relations

**Definition ($h$-to-$h$-routing problem)**

An $h$-relation is a routing problem in which every node is the source of $h$ packets and the destination of $h$ packets.

- Observe that a "1-relation" is a permutation routing problem."
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**Lemma**

*Suppose we use Valiant’s trick for routing an arbitrary $h$-relation on the hypercube. The congestion $C$ is $O(\log n + h)$, w.h.p.*

Proof: Exercise
Scheduling on the hypercube

We study the problem of forwarding packets along prespecified paths on the $d$-dimensional hypercube.

Theorem

Suppose we are given a set of packets each of which coming with a bit-fixing path along which it should be sent from its source to its destination.
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There is a distributed, randomized scheduling protocol that delivers all packets in time $O(C + \log n)$, w.h.p.
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There is a distributed, randomized scheduling protocol that delivers all packets in time \(O(C + \log n)\), w.h.p.

Combining this result with Valiant’s trick gives:

**Corollary**

There is a distributed algorithm that routes any \(h\)-relation in time \(O(h + \log n)\), w.h.p., on the hypercube.
Randomized scheduling policy

The random rank protocol:

- Let $R$ denote a sufficiently large integer whose value will be specified later.
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- If two or more packets contend for the same edge in a step, then the one with smallest rank is forwarded and the others have to wait.
- In case of equal ranks, packets with smaller ids are preferred.
Delay sequence analysis

Our analysis uses the following witness structure.

Definition (delay sequence)

A delay sequence $DS$ of length $s$ consists of

1. a delay path $P = (e(1), \ldots, e(L))$, $1 \leq L \leq d$, with edges of increasing dimension (like a bit-fixing path in reverse order)
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**Definition (active delay sequence)**

$DS$ is called active if $r(p_i) = k_i$, for $0 \leq i \leq s$. 
Delay sequence analysis

Lemma

If the random rank protocol needs $T > d$ steps, then there exists an active DS of length at least $T - d$.

Proof:

- Consider any packet $p$ arriving at its destination in step $T$. As $T > d$, this packet must have been delayed for at least one step. We call this packet $p_0$. 
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- Next we follow packet $p_2$ and so on until we reach a packet $p_s$, $s \geq 1$, that was not delayed before. We follow this packet back to its source.

- Our tour backward through time covers $T$ steps and we observed $s$ delays. Let $L$ denote the number of edges on the recorded path.
Delay sequence analysis

- From this tour backwards through time, we can now construct an active DS as follows.
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5. For \( 0 \leq i \leq s \), we set \( k_i = r(p_i) \). Observe that this gives \( k_0 \geq k_1 \geq \cdots \geq k_s \) as packet \( p_{i-1} \) is delayed by packet \( p_i \) and the protocol prefers packets with smaller rank.
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This ends the proof of the lemma.
Delay sequence analysis

Now we bound the probability that there exists an active DS. Our analysis begins with counting delay sequences.

Lemma

The number of delay sequences of length \( s \) is at most

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n^2 \cdot \binom{L - 1 + s}{s} \cdot C^{s+1} \cdot \binom{R + s}{s + 1}.
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Delay sequence analysis

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Lemma

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Proof:

1) Counting delay paths:
The number of ways to choose a delay path is $n(n - 1) \leq n^2$ as this path corresponds to a bit-fixing path (in reverse order) that is determined by specifying the first and the last node on the path.
Delay sequence analysis

The number of delay sequences of length $s$ is at most $n^2 \cdot \binom{L-1+s}{s} \cdot C^{s+1} \cdot \binom{R+s}{s+1}$.

2) Counting the ways to choose the $\ell_i$’s and the $k_i$’s:
How many ways are there to choose the integers $\ell_1, \ldots, \ell_s$ such that $1 \leq \ell_1 \leq \ell_2 \leq \cdots \leq \ell_s \leq d$?
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These integers can be encoded into a binary string as follows

$$0^{\ell_1-1}10^{\ell_2-\ell_1}10^{\ell_3-\ell_2}1\ldots10^{\ell_s-\ell_{s-1}}10^{d-\ell_s}.$$
Delay sequence analysis

The number of delay sequences of length $s$ is at most $n^2 \cdot \binom{s+1}{R+s} \cdot \binom{s}{L-s+1} \cdot s\cdot (d-s+1)$.

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$$0^{\ell_1-1}10^{\ell_2-\ell_1}10^{\ell_3-\ell_2}1\ldots10^{\ell_s-\ell_{s-1}}10^{d-\ell_s}.$$

Observe that this string contains $s$ ones and the number of zeros in this string is

$$\ell_1 - 1 + \left( \sum_{i=2}^{s} (\ell_i - \ell_{i-1}) \right) + d - \ell_s = d - 1.$$
Delay sequence analysis

The number of delay sequences of length $s$ is at most $n^2 \cdot \binom{L-1+s}{s} \cdot C^{s+1} \cdot \binom{R+s}{s+1}$.

Consequently, there is a one-to-one mapping between the $\ell_i$'s and the binary strings with $d - 1$ zeros and $s$ ones. Hence, the number of ways to choose the $\ell_i$'s corresponds to the number of such strings which is

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- Analogously the number of ways to choose $k_0, \ldots, k_s \in [R]$ such that $k_0 \geq k_1 \geq \cdots \geq k_s$ is equal to the number of binary strings consisting of $R - 1$ zeroes and $s + 1$ ones, which is

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The number of delay sequences of length $s$ is at most $n^2 \cdot \binom{L-1+s}{s} \cdot C^{s+1} \cdot \binom{R+s}{s+1}$.

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Now suppose that the delay path $P$ and the $\ell_i$'s are fixed.
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Now suppose that the delay path $P$ and the $\ell_i$'s are fixed.

Then, for each delay packet, we know an edge that is contained in its path: In particular, we know that packet $p_i$, for $1 \leq i \leq s$, uses edge $e(\ell_i)$ and packet $p_0$ uses edge $e(\ell_1)$. 

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How many possibilities are there to choose a packet whose path is leading through a known edge? – At most $C$ since each edge is contained in the paths of at most $C$ packets.
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Hence, there are at most $C$ possibilities to choose $p_i$ and, hence, at most $C^{s+1}$ possibilities to choose all delay packets $p_0, \ldots, p_s$. 
Delay sequence analysis

The number of delay sequences of length $s$ is at most $n^2 \cdot \left( \frac{L^{-1+s}}{s} \right) \cdot C^{s+1} \cdot \left( \frac{R+s}{s+1} \right)$.

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Lemma

The probability that a given DS of length $s$ is active is $R^{-(s+1)}$.

Proof:

- For every delay packet $p_i$, the probability that the packet’s rank is $k_i$ is $1/R$ because ranks are chosen uniformly at random from $[R]$. 
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- By the first Lemma, if the algorithm needs $T \geq d + s$ steps, then there exists an active delay sequence of length at least $s$.
- Cutting this sequence after packet $p_s$ gives an active delay sequence of length exactly $s$.
- Let $\mathcal{DS}(s)$ denote the set of delay sequences of length $s$. It holds

$$
\Pr[T \geq d + s] \leq \Pr[\exists DS \in \mathcal{DS}(s) : DS \text{ is active}]
\leq \sum_{DS \in \mathcal{DS}(s)} \Pr[DS \text{ is active}]
= \sum_{DS \in \mathcal{DS}(s)} R^{-(s+1)}
\leq n^2 \cdot \binom{d - 1 + s}{s} \cdot C^{s+1} \cdot \binom{R + s}{s + 1} \cdot R^{-(s+1)}.
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Delay sequence analysis

We have so far:

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- Using \((\frac{a}{b}) \leq 2^a\) and \((\frac{a}{b}) \leq \left( \frac{ea}{b} \right)^b\) to upper-bound the binomial coefficients, we obtain

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\Pr[T \geq d + s] \leq n^2 \cdot 2^{d-1+s} \cdot C^{s+1} \cdot \left( \frac{e(R + s)}{s + 1} \right)^{s+1} \cdot R^{-(s+1)}
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\[ \leq n^3 \cdot \left(\frac{2eC(R + s)}{(s + 1)R}\right)^{s+1} \]

- Choosing \( R \geq s \) yields \( R + s \leq 2R \) and, hence,

\[ \Pr[T \geq d + s] \leq n^3 \cdot \left(\frac{4eC}{s + 1}\right)^{s+1} \]
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- Hence, with probability at least \( 1 - n^{-\alpha} \), the random rank protocol needs at most \( d + s - 1 = O(C + \log n) \) steps.
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- Hence, with probability at least \( 1 - n^{-\alpha} \), the random rank protocol needs at most \( d + s - 1 = O(C + \log n) \) steps.

- This end the proof of the theorem.
Consider any $n$-node network $G = (V, E)$.\textsuperscript{4} We study the problem of forwarding packets along arbitrary shortest paths in $G$.

**Theorem**

Suppose we are given a set of $N \geq n$ packets each of which coming with a shortest path in $G$ along which it should be sent from its source to its destination.

Let $C$ and $D$ denote the congestion and the dilation of the paths, respectively. There is a distributed, randomized scheduling protocol that delivers all packets in time $O(C + D + \log N)$, w.h.p.

\textsuperscript{4}Recall that edges are assumed to be directed. In order to represent an undirected network, one replaces each edge by two directed edges in opposite direction.
Randomized scheduling policy with increasing ranks

The growing rank protocol:

- Let $R$ denote a sufficiently large integer being a multiple of $D$. 
Randomized scheduling policy with increasing ranks

The growing rank protocol:

- Let \( R \) denote a sufficiently large integer being a multiple of \( D \).
- Every packet \( p \) is assigned independently and uniformly at random a rank \( r(p) \in [R] \).
Randomized scheduling policy with increasing ranks

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1. Let $R$ denote a sufficiently large integer being a multiple of $D$.
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- Besides every packet is assumed to have a unique integer ident number (id).
- If two or more packets contend for the same edge in a step, then the one with smallest rank is forwarded and the others have to wait.
- In case of equal ranks, packets with smaller ids are preferred.
Observation

As the initial rank is at most $R - 1$ and the rank of a packet is increased at most $D$ times by $R/D$, the final rank of a packet is at most $2R - 1$.

Let $r_e(p) \in [2R]$ denote the rank of packet $p$ in those time steps in which $p$ contends for being forwarded along edge $e$. 
Delay sequence analysis

We adapt the definition of a delay sequence as follows.

**Definition (delay sequence)**

A delay sequence $DS$ of length $s$ consists of

1. a delay path $P = (e(1), \ldots, e(L))$, for $L \leq 2D$, with edges in reverse direction, that is, $(e(L), \ldots, e(1))$ is a path in $G$;
2. $s$ numbers $\ell_1, \ldots, \ell_s \in \{1, \ldots, L\}$ with $\ell_1 \leq \ell_2 \leq \cdots \leq \ell_s$;
3. $s + 1$ distinct delay packets $p_0, p_1, \ldots, p_s$ such that, for $1 \leq i \leq s$, edge $e(\ell_i)$ is contained in the paths of packet $p_{i-1}$ and packet $p_i$;
4. $s + 1$ numbers $k_0, k_1, \ldots, k_s \in [2R]$ with $k_0 \geq k_1 \geq \cdots \geq k_s$.

**Definition (active delay sequence)**

$DS$ is active if $r_e(\ell_i)(p_i) = k_i$, for $1 \leq i \leq s$, and $r_e(\ell_1)(p_0) = k_0$. 
Lemma

If the growing rank protocol needs $T \geq 2D$ steps, then there exists an active DS of length at least $T - 2D$.

Proof:
Consider any packet packet arriving at its destination in step $T$. As $T \geq 2D$, this packet must have been delayed for at least one step. We call this packet $p_0$. We follow the path of $p_0$ backwards through time from its destination until we reach an edge where it has been delayed by a packet that we call $p_1$. Now we follow $p_1$ backwards through time until we reach a time step where this packet has been delayed before by another packet that we call $p_2$, and so on ... ... until we reach a packet $p_s$, for some $s \geq 1$, that was not delayed before. We follow this packet back to its source.
Delay sequence analysis

From this tour backwards through time, we can now construct an active $DS$ as follows.

- The path that we have recorded by this process in reverse order gives us the delay path $P = (e(1), \ldots, e(L))$.

This is the only part of the analysis where we need to assume that the paths of the packets are shortest paths in $G$.\(^5\)
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- For \( 1 \leq i \leq s \), we set \( k_i = r_{e(\ell_i)}(p_i) \), and \( k_0 = r_{e(\ell_1)}(p_0) \).

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- The packets \( p_0, \ldots, p_s \) are defined to be the delay packets.
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- For \( 1 \leq i \leq s \), we set \( k_i = r_{e(\ell_i)}(p_i) \), and \( k_0 = r_{e(\ell_1)}(p_0) \).

**Exercise:** Show that the packets \( p_0, \ldots, p_s \) are distinct, that is, no packet appears more than once in the delay sequence.\(^5\)

---

\(^5\)This is the only part of the analysis where we need to assume that the paths of the packets are shortest paths in \( G \).
Delay sequence analysis

Observe that $k_0 \geq k_1 \geq \cdots \geq k_s$ as the ranks of the delay packets do not increase on our tour. More specifically:

- whenever we switch from packet $p_i$ to packet $p_{i+1}$ on our tour, the rank of $p_{i+1}$ is not larger than the rank of packet $p_i$ because $p_{i+1}$ delays $p_i$ and the protocol prefers packets with smaller rank, and
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- whenever we add an edge to the delay path and follow this edge, the rank of the currently observed packet is decreased (by \( R/d \)) as we follow the packet backwards in time.
Delay sequence analysis

It only remains to prove \( L \leq 2D \) and \( s \geq T - 2D \).

- The final rank of \( p_0 \) is at most \( 2R - 1 \).
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- As ranks are non-negative, we obtain $(2R - 1) - L(R/D) \geq 0$ which gives $L \leq (2R - 1)/(R/D) \leq 2D$. 

Finally, $T = L + s$ implies $s = T - L \geq T - 2D$. 

This ends the proof of the lemma.
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Delay sequence analysis

Now we bound the probability that there exists an active DS. Our analysis begins with counting delay sequences.

Lemma

The number of delay sequences of length $s$ is at most

$$\binom{2D-1+s}{s} \binom{2R+s}{s+1} N C^s.$$
Delay sequence analysis

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**Lemma**

*The number of delay sequences of length $s$ is at most*

$$\left( \frac{2D - 1 + s}{s} \right) \left( \frac{2R + s}{s + 1} \right) N C^s.$$

**Proof:**
Analogously to the analysis for the hypercube the number of ways to choose the $\ell_i$’s and the $k_i'$ can be bounded by

$$\left( \frac{2D - 1 + s}{s} \right) \left( \frac{2R + s}{s + 1} \right).$$
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Now we assume that the $\ell_i$’s are fixed and we count the number of ways to choose the delay packets and the edges on the delay path:

- There are $N$ possibilities to choose packet $p_0$. 
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- There are $N$ possibilities to choose packet $p_0$.
- Once $p_0$ is fixed, we can construct the first part of the delay path from edge $e(1)$ up edge $e(\ell_1)$ by following the path of $p_0$ backwards from its destination.
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- Now, as the path of $p_1$ contains $e(\ell_1)$, there are at most $C$ possibilities to choose $p_1$.
- Once $p_1$ is fixed, we can determine the delay path up to $e(\ell_2)$. 
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- Once $p_1$ is fixed, we can determine the delay path up to $e(\ell_2)$.

- As the path of $p_2$ contains $e(\ell_2)$, there are again at most $C$ possibilities to choose $p_2$, and so on until packet $p_5$. 
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- Now, as the path of $p_1$ contains $e(\ell_1)$, there are at most $C$ possibilities to choose $p_1$.
- Once $p_1$ is fixed, we can determine the delay path up to $e(\ell_2)$.
- As the path of $p_2$ contains $e(\ell_2)$, there are again at most $C$ possibilities to choose $p_2$, and so on until packet $p_s$.
- Thus, the number of possibilities to choose the delay packets and to construct the delay path is at most $NC^s$. 
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- Once $p_0$ is fixed, we can construct the first part of the delay path from edge $e(1)$ up edge $e(\ell_1)$ by following the path of $p_0$ backwards from its destination.
- Now, as the path of $p_1$ contains $e(\ell_1)$, there are at most $C$ possibilities to choose $p_1$.
- Once $p_1$ is fixed, we can determine the delay path up to $e(\ell_2)$.
- As the path of $p_2$ contains $e(\ell_2)$, there are again at most $C$ possibilities to choose $p_2$, and so on until packet $p_s$.
- Thus, the number of possibilities to choose the delay packets and to construct the delay path is at most $NC^s$.
- This ends the proof of the lemma.
Delay sequence analysis

**Lemma**

The probability that a DS of length $s$ is active is at most $R^{-(s+1)}$.

**Proof:**

- Suppose $e(\ell_i)$ is the $j$th edge on the path of packet $p_i$. 

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- Suppose $e(\ell_i)$ is the $j$th edge on the path of packet $p_i$.
- The rank at edge $e(\ell_i)$ is equal to $k_i$ if its initial rank is equal to $k'_i = k_i - (j - 1) \cdot R/D$, which happens with probability $1/R$ if $k'_i \in [R]$, and probability 0, otherwise.
Lemma

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- That is, the probability that the rank of \( p_i \) at edge \( e(\ell_i) \) is equal to \( k_i \) is at most \( 1/R \).
Delay sequence analysis

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The probability that a DS of length s is active is at most $R^{-(s+1)}$.

Proof:

- Suppose $e(l_i)$ is the $j$th edge on the path of packet $p_i$.
- The rank at edge $e(l_i)$ is equal to $k_i$ if its initial rank is equal to $k'_i = k_i - (j - 1) \cdot R/D$, which happens with probability $1/R$ if $k'_i \in [R]$, and probability 0, otherwise.
- That is, the probability that the rank of $p_i$ at edge $e(l_i)$ is equal to $k_i$ is at most $1/R$.
- Consequently, the probability that all $s + 1$ delay packets have the prescribed rank is at most $R^{-(s+1)}$. 
Delay sequence analysis

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- Consequently, the probability that all $s + 1$ delay packets have the prescribed rank is at most $R^{-(s+1)}$.
- This ends the proof of the lemma.
Delay sequence analysis

Now we proceed analogously to the analysis for the hypercube.

\[
\Pr[T \geq 2D + s] \leq \Pr[\exists DS \in DS(s) : DS \text{ is active}]
\leq \sum_{DS \in DS(s)} \Pr[DS \text{ is active}]
\leq \binom{2D - 1 + s}{s} \binom{2R + s}{s + 1} N C^s R^{-(s+1)}
\leq 2^{2D-1+s} \left(\frac{e(2R + s)}{s + 1}\right)^{s+1} N C^s R^{-(s+1)}
\leq 2^{2D} \left(\frac{6Ce}{s + 1}\right)^{s+1} N ,
\]

where the last equation assumes \( R \geq s \).
Finally, we set $s = \lceil \max\{12eC, (\alpha + 1) \log N + 2D\} \rceil - 1 = O(C + D + \log N)$. This gives

$$\Pr[T \geq 2D + s] \leq 2^{2D} N \left(\frac{1}{2}\right)^{s+1} \leq 2^{2D} N \left(\frac{1}{2}\right)^{(\alpha+1) \log N+2D} \leq N^{-\alpha} \leq n^{-\alpha}$$

using $n \leq N$.

Hence, with probability at least $1 - n^{-\alpha}$, the growing rank protocol needs at most $2D + s - 1 = O(C + D + \log N)$ steps.
Literature