

Improved Distance Sensitivity Oracles with Subcubic Preprocessing Time

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Abstract

We consider the problem of building *Distance Sensitivity Oracles* (DSOs). Given a directed graph $G = (V, E)$ with edge weights in $\{1, 2, \dots, M\}$, we need to preprocess it into a data structure, and answer the following queries: given vertices $u, v \in V$ and a failed vertex or edge $f \in (V \cup E)$, output the length of the shortest path from u to v that does not go through f . Our main result is a simple DSO with $\tilde{O}(n^{2.7233}M)$ preprocessing time and $O(1)$ query time. Moreover, if the input graph is undirected, the preprocessing time can be improved to $\tilde{O}(n^{2.6865}M)$. The preprocessing algorithm is randomized with correct probability $\geq 1 - 1/n^C$, for a constant C that can be made arbitrarily large. Previously, there is a DSO with $\tilde{O}(n^{2.8729}M)$ preprocessing time and $\text{polylog}(n)$ query time [Chechik and Cohen, STOC'20].

At the core of our DSO is the following observation from [Bernstein and Karger, STOC'09]: if there is a DSO with preprocessing time P and query time Q , then we can construct a DSO with preprocessing time $P + \tilde{O}(n^2) \cdot Q$ and query time $O(1)$. (Here $\tilde{O}(\cdot)$ hides $\text{polylog}(n)$ factors.)

2012 ACM Subject Classification Theory of computation \rightarrow Shortest paths; Theory of computation \rightarrow Data structures design and analysis

Keywords and phrases Graph theory, Shortest paths, Distance sensitivity oracles

Related Version There is a paper with the same title accepted to ESA'20 [17]. Compared to that version, here we use a new argument dealing with non-unique shortest paths, thus improve the preprocessing time from $\tilde{O}(n^{2.7233}M^2)$ to $\tilde{O}(n^{2.7233}M)$. We also extend our argument to edge failures, instead of only vertex failures as in [17].

Acknowledgements I would like to thank Zihao Li for introducing this problem to me, Ran Duan and Yong Gu for helpful discussions, and Ce Jin for helpful comments that improve the presentation of this paper. I would also like to thank Hongxun Wu for reading and commenting on an early draft version of this paper, and pointing out the issue with non-unique shortest paths. I am also grateful to the anonymous referees of ESA for their helpful comments that improve the presentation of this paper, and for suggesting me to consider the more general case that edge weights may be negative. (Unfortunately, I am not aware of any easy way to generalize our results to handle negative edge weights.)

1 Introduction

Suppose we are given a directed graph $G = (V, E)$, and we want to build a data structure that, given any two vertices $u, v \in V$ and a failure f , which is either a failed vertex or a failed edge, outputs the length of the shortest path from u to v that does not go through f . Such a data structure is called a *Distance Sensitivity Oracle* (or DSO for short).

The problem of constructing DSOs is motivated by the fact that real-life networks often suffer from failures. Suppose we have a network with n nodes and m (directed) links, and we want to send a package from a node u to another node v . Normally, it suffices to compute the shortest path from u to v . However, if some node or link f in this network fails, then our path cannot use f , and our task becomes to find the shortest path from u to v that does not go through f . Usually, there is only a very small number of failures. In this paper, we consider the simplest case, in which there is only one failed node or link.

The problem of constructing a DSO is well-studied: Demetrescu et al. [6] showed that given a directed graph $G = (V, E)$, there is a DSO which occupies $O(n^2 \log n)$ space, and can answer a query in $O(1)$ time. Duan and Zhang [9] improved the space complexity to $O(n^2)$, which is optimal for dense graphs (i.e. $m = \Theta(n^2)$).

Unfortunately, the oracle in [6] requires a large preprocessing time ($O(mn^2 + n^3 \log n)$). In real-life applications, the preprocessing time of the DSO is also very important. Bernstein and Karger [2, 3] improved this time bound to $\tilde{O}(mn)$. Note that the All-Pairs Shortest Paths (APSP) problem, which only asks the distances between each pair of vertices u, v , is conjectured to require $mn^{1-o(1)}$ time to solve [14]. Since we can solve the APSP problem by using a DSO, the preprocessing time $\tilde{O}(mn)$ is optimal in this sense.

However, if the edge weights are small positive integers (that do not exceed M), then the APSP problem can be solved in $\tilde{O}(n^{2.5286}M)$ time [22]¹, which is significantly faster than $O(mn)$ for dense graphs with small weights (e.g. $M = O(1)$). Thus it might be possible to obtain better results than [3] in the regime of small integer edge weights. Weimann and Yuster [21] showed that for any constant $\alpha \in (0, 1)$, we can construct a DSO in $\tilde{O}(n^{1-\alpha+\omega}M)$ time. Here $\omega < 2.3728639$ is the exponent of matrix multiplication [11]. However, the query time for this oracle is $\tilde{O}(n^{1+\alpha})$, which is *superlinear*. Later, Grandoni and Williams [13] showed that for every constant $\alpha \in [0, 1]$, we can construct a DSO in $\tilde{O}(n^{\omega+1/2}M + n^{\omega+\alpha(4-\omega)}M)$ time, which answers each query in $\tilde{O}(n^{1-\alpha})$ time.

Recently, in an independent work, Chechik and Cohen [4] showed that a DSO with $\text{polylog}(n)$ query time can be constructed in $\tilde{O}(Mn^{2.873})$ time, achieving *both* subcubic preprocessing time and polylogarithmic query time. The space complexity for their DSO is $\tilde{O}(n^{2.5})$.

1.1 Our Results

In this work, we show improved and simplified constructions of DSOs. We start with an observation.

► **Observation 1.1** (informal). *If we have a DSO with preprocessing time P and query time Q , then we can build a DSO with preprocessing time $P + \tilde{O}(n^2) \cdot Q$ and query time $O(1)$.*

For $\alpha = 0.2$, the oracle in [13] already achieves $\tilde{O}(n^{2.8729}M)$ preprocessing time and $O(n^{0.8})$ query time. Observation 1.1 implies that this query time can be brought down to $O(1)$.

Observation 1.1 can be proved by a close inspection of [3]: The algorithm in [3] for constructing a DSO picks $\tilde{O}(n^2)$ carefully chosen queries (u, v, x) , such that the answers of all these queries can be computed in $\tilde{O}(mn)$ time. Then from these answers, we can easily compute a DSO with $O(1)$ query time. If, instead of computing these answers in $\tilde{O}(mn)$ time, we use the given DSO to answer these queries, the preprocessing time becomes $P + \tilde{O}(n^2) \cdot Q$.

Our main result is a simple construction of DSOs with preprocessing time $\tilde{O}(n^{2.7233}M)$ and query time $O(1)$. If the input graph is undirected, we can achieve a better preprocessing time of $\tilde{O}(n^{2.6865}M)$.

► **Theorem 1.2.** *We can construct a DSO with $\tilde{O}(n^{2.7233}M)$ preprocessing time and $O(1)$ query time. Moreover, if the input graph is undirected, then we can construct a DSO with*

¹ [22] only claims a time bound of $\tilde{O}(n^{2.58}M)$; the time bound $\tilde{O}(n^{2.5286}M)$ is calculated using improved time bounds for rectangular matrix multiplication, see [12].

$\tilde{O}(n^{(3+\omega)/2}M) = \tilde{O}(n^{2.6865}M)$ preprocessing time and $O(1)$ query time. The construction algorithms are randomized and yield valid DSOs w.h.p.²

Compared to previous constructions of DSOs, our results have better dependence on n , and is conceptually simpler. Also, the space complexity of our DSO is $\tilde{O}(n^2)$, which is better than [4]. However, one drawback of our construction is that it cannot handle negative edge weights. (The preprocessing and query time bounds stated above for [4, 13, 21] still hold when the edge weights are integers in $[-M, M]$.)

1.2 Non-unique Shortest Paths

A subtle technical issue is that the shortest paths in G between some vertex pairs may not be unique. Normally, if we perturb every edge weight by a small random value, then we can ensure that shortest paths are unique w.h.p. by the isolation lemma [16, 20]. However, the subcubic-time algorithms for APSP [18, 19, 22] depend crucially on the assumption that edge weights are small integers, so we cannot perform the random perturbation.

Fortunately, there is another elegant way of breaking ties, described in Section 3.4 of [5], that is suitable for our propose. We arbitrarily assign a linear order \prec over the edges of G . (For example, we index the edges by distinct numbers in $\{1, 2, \dots, m\}$, and define $e_1 \prec e_2$ if the index of e_1 is smaller than the index of e_2 .) For a path p , define the *bottleneck edge* of p as the smallest edge (w.r.t. \prec) in p . Let G' be a subgraph of G (think of $G' = G$ or G' is G with some vertex or edge removed), define $w_{G'}(u, v)$ as the largest (w.r.t. \prec) edge $w \in G'$ such that there is a *shortest* path in G' from u to v whose bottleneck edge is w . Define $\rho_{G'}(u, v)$ to be the following path: If $u = v$, then $\rho_{G'}(u, v)$ is the empty path; otherwise, suppose $w_{G'}(u, v)$ is an edge from u^* to v^* , then $\rho_{G'}(u, v)$ is the concatenation of $\rho_{G'}(u, u^*)$, $w_{G'}(u, v)$ and $\rho_{G'}(v^*, v)$. Note that $\rho_{G'}(u, v)$ is always well-defined: we can define $\rho_{G'}(u, v)$ inductively, in non-decreasing order of the distance from u to v .

We list a few properties of this tie-breaking method in the following theorem.

- **Theorem 1.3.** *The following properties of $\rho_{G'}(u, v)$ are true:*
- (Property a) *For every $u', v' \in \rho_{G'}(u, v)$ such that u' appears before v' , the portion of $u' \rightsquigarrow v'$ in $\rho_{G'}(u, v)$ coincides with the path $\rho_{G'}(u', v')$.*
- (Property b) *Let G' be a subgraph of G , suppose $\rho_G(u, v)$ is completely contained in G' , then $\rho_{G'}(u, v) = \rho_G(u, v)$.*

For every vertex $u \in V$, we can compute the collection of paths $\rho_{G'}(v, u)$ for every $v \in V$. By (Property a), these paths form a tree rooted at u , and we say this tree is the *incoming shortest path tree* rooted at u , denoted as $T^{\text{in}}(u)$. Similarly we can define the outgoing shortest path trees rooted at each vertex, denoted as $T^{\text{out}}(u)$. The advantage of this tie-breaking method is:

- **Theorem 1.4** ([8]). *Given a subgraph G' , the incoming and outgoing shortest path trees $T^{\text{in}}(u)$ and $T^{\text{out}}(u)$ for each vertex u can be computed in $\tilde{O}(n^{(3+\omega)/2}M^{1/2})$ time.*

For completeness we will prove Theorem 1.3 and 1.4 in Appendix A.

1.3 Notation

We mainly stick to the notation used in [7], namely:

² We say that an event happens *with high probability* (w.h.p.), if it happens with probability $1 - 1/n^C$, for some constant C that can be made arbitrarily large.

- If p is a path, then $|p|$ denotes the number of edges in it, and $\|p\|$ denotes its length (i.e. the total weight of its edges).
- Let $u, v \in V$, we denote $uv = \rho_G(u, v)$, i.e. the unique shortest path from u to v in the original graph. For a vertex or edge failure f , we denote $uv \diamond f = \rho_{G-f}(u, v)$, where $G - f$ is the subgraph of G obtained by removing the failure f . Note that if f is not in the original path uv , then by (Property b), $uv \diamond f = uv$.
- Let p be a path from u to v . For two vertices $a, b \in p$ such that a appears earlier than b , we say the interval $p[a, b]$ is the subpath from a to b , and $p(a, b)$ is the path $p[a, b]$ without its endpoints (a and b). If the path p is known in the context, then we may omit p and simply write $[a, b]$ or (a, b) . (Property a) implies that, if p is the path uv (or $uv \diamond f$ resp.), then $p[a, b]$ is the path ab (or $ab \diamond f$ resp).

We define $\text{MM}(n_1, n_2, n_3)$ as the complexity of multiplying an $n_1 \times n_2$ matrix and an $n_2 \times n_3$ matrix. Let a, b, c be real numbers, we define $\omega(a, b, c)$ be the infimum over all real numbers α such that $\text{MM}(n^a, n^b, n^c) = O(n^\alpha)$. For any real number r , we have $\omega(1, 1, r) = \omega(1, r, 1) = \omega(r, 1, 1)$ [15], and we denote $\omega(r) = \omega(1, 1, r)$.

We also need the following adaptation of Zwick’s APSP algorithm [22] (see also [13, Corollary 3.1]):

► **Theorem 1.5.** *Given a directed graph $G = (V, E)$ with edge weights in $\{1, 2, \dots, M\}$, and an integer r , we can compute the distances between every pair of nodes whose shortest path uses at most r edges, in $\tilde{O}(rM \cdot \text{MM}(n, n/r, n))$ time.*

Proof Sketch. We simply run the first $\lceil \log_{3/2} r \rceil$ iterations of the algorithm **rand-short-path** in [22]. The correctness of this algorithm is guaranteed by [22, Lemma 4.2]. ◀

2 Constructing a DSO in $\tilde{O}(n^{2.7233}M)$ Time

In this section, we prove Theorem 1.2.

► **Theorem 1.2.** *We can construct a DSO with $\tilde{O}(n^{2.7233}M)$ preprocessing time and $O(1)$ query time. Moreover, if the input graph is undirected, then we can construct a DSO with $\tilde{O}(n^{(3+\omega)/2}M) = \tilde{O}(n^{2.6865}M)$ preprocessing time and $O(1)$ query time. The construction algorithms are randomized and yield valid DSOs w.h.p.*

Given an integer r and a graph $G = (V, E)$, we define an r -truncated DSO to be a data structure that, when given a query (u, v, f) , outputs the value $\min\{\|uv \diamond f\|, r\}$. In other words, an r -truncated DSO is a DSO which only needs to be correct when the path $uv \diamond f$ has length at most r . If this length is greater than r , it outputs r instead.

Inspired by Zwick’s APSP algorithm [22], our main idea is to compute an r -truncated DSO for every $r = (3/2)^i$. Our strategies for small r and large r are completely different.

When r is small, the sampling approach in [13, 21] already suffices. Fix a particular query (u, v, f) , we assume that f is a vertex failure, and $\|uv \diamond f\| \leq r$. In particular, since the edge weights are positive, there are at most $r + 1$ vertices in the path $uv \diamond f$. Suppose we sample a graph by discarding each vertex w.p. $1/r$. With probability $\Omega(1/r)$, the resulting graph would “capture” this query in the sense that f is not in it but $uv \diamond f$ is completely included in it. Therefore, if we take $\tilde{O}(r)$ independent samples, and compute APSP for each sampled subgraph, we can deal with every vertex-failure query w.h.p. If f is an edge failure, we simply discard every *edge* (instead of vertex) in every subgraph w.p. $1/r$, and we can still deal with every edge-failure query w.h.p.

For large r , our idea is to compute a $(3/2)r$ -truncated DSO from an r -truncated DSO. More precisely, given an r -truncated DSO with $O(1)$ query time, we can compute a $(3/2)r$ -truncated DSO with $\tilde{O}(Mn/r)$ query time. First we sample a bridging set (see [22]) H of size $\tilde{O}(Mn/r)$. Let (u, v, f) be a query such that $r \leq \|uv \diamond f\| \leq (3/2)r$, then w.h.p. there is a “bridging vertex” $h \in H$ such that h is on the path $uv \diamond f$, and both of the queries (u, h, f) and (h, v, f) are captured by the r -truncated DSO. If we iterate through H , we can answer the query (u, v, f) in $\tilde{O}(Mn/r)$ time. Then we use an “ r -truncated” version of Observation 1.1 to transform this $(3/2)r$ -truncated DSO into a new one with $O(1)$ query time.

2.1 Case I: r is Small

Let $\tilde{r} = \lceil 8Cr \ln n \rceil$, where C is a large enough constant. We independently sample $2\tilde{r}$ graphs $G_1^e, G_2^e, \dots, G_{\tilde{r}}^e$ and $G_1^v, G_2^v, \dots, G_{\tilde{r}}^v$. (The superscripts e and v stand for “edge” and “vertex” respectively.) Each graph is sampled as follows:

- The vertex set of each G_i^e is equal to the original vertex set V , and the edge set of each G_i^e is sampled by including every edge independently w.p. $1 - 1/r$.
- The vertex set of each G_i^v is sampled by including every vertex independently w.p. $1 - 1/r$, and the graph G_i^v is the induced subgraph of G on vertices $V(G_i^v)$.

Then, for each $1 \leq i \leq \tilde{r}$, we compute all-pairs shortest paths of the graph G_i^e and G_i^v , but we only compute the shortest paths that use at most r edges. By Theorem 1.5, this step can be done in $\tilde{O}(rM \cdot \text{MM}(n, n/r, n))$ time for each graph. Alternatively, if the input graph is undirected, then this step can be done in $\tilde{O}(Mn^\omega)$ time [18, 19] for each graph.

Consider a query (u, v, f) , suppose f is a vertex failure, and assume that $\|uv \diamond f\| \leq r$. Let $1 \leq i \leq \tilde{r}$, we say i is *good* for the query (u, v, f) , if both of the following hold.

- The graph G_i^v does not contain f .
- The graph G_i^e contains the entire path $uv \diamond f$.

For every i ($1 \leq i \leq \tilde{r}$), the probability that i is good for the particular query (u, v, f) is at least

$$(1/r) \cdot (1 - 1/r)^{r+1} \geq 1/8r \text{ (if } r \geq 2\text{)}.$$

Given a query (u, v, f) , where f is a vertex failure, we iterate through all i 's such that $f \notin V(G_i^v)$, and take the smallest value among the distances from u to v in these graphs G_i^e . With high probability, there are only $\tilde{O}(1)$ valid i 's such that $f \notin V(G_i^v)$, and we can preprocess this set of i 's for every $f \in V$. Therefore the query time is $\tilde{O}(1)$.

The algorithm succeeds at a query (u, v, f) if there is an i that is good for (u, v, f) . Since the G_i 's are independent, the probability that there is an i good for (u, v, f) is at least

$$1 - (1 - 1/8r)^{\tilde{r}} \geq 1 - 1/n^C.$$

By a union bound over all n^3 triples of possible queries (u, v, f) , it follows that our data structure is correct w.p. at least $1 - 1/n^{C-3}$, which is a high probability.

If f is an edge failure, we look at the graphs G_i^e instead of G_i^v . We say i is good if the graph G_i^e does not contain f but contains the entire path $uv \diamond f$. Again, if we fix (u, v, f) , then the probability that a particular i is good is at least $1/\Omega(r)$, so w.h.p. there is some i that is good for (u, v, f) . By a union bound over all $O(n^4)$ triples (u, v, f) , our data structure is still correct w.h.p.

In conclusion, there is an r -truncated DSO with $\tilde{O}(1)$ query time, whose preprocessing time is $\tilde{O}(\tilde{r} \cdot rM \cdot \text{MM}(n, n/r, n))$ for directed graphs, and $\tilde{O}(\tilde{r} \cdot Mn^\omega)$ for undirected graphs.

2.2 An Observation

We need the following observation (“ r -truncated” version of Observation 1.1), which roughly states that given an r -truncated DSO with preprocessing time P and query time Q , we can build an r -truncated DSO with preprocessing time $P + \tilde{O}(n^2) \cdot Q$ and query time $O(1)$. More formally, we have:

► **Observation 2.1.** *Let r be an integer, $G = (V, E)$ be an input graph, and \mathcal{D} be an arbitrary r -truncated DSO. We can construct $\text{Fast}(\mathcal{D})$, which is an r -truncated DSO with $O(1)$ query time and a preprocessing algorithm as follows.*

- *It first computes the distance matrix of G , and the (incoming and outgoing) shortest path trees rooted at each vertex.*
- *Then it invokes the preprocessing algorithm of \mathcal{D} on the input graph G .*
- *At last, it makes $\tilde{O}(n^2)$ queries to \mathcal{D} , and spends $\tilde{O}(n^2)$ additional time to finish the preprocessing algorithm.*

We emphasize the following technical details that are not reflected in the informal statement of Observation 1.1. First, we need to compute the distance matrix and shortest path trees of G (henceforth the “APSP data” of G) before using \mathcal{D} . We can compute the APSP data using Theorem 1.4. Second, the preprocessing algorithm of \mathcal{D} is called *only once*, and on the same graph G (on which we have already computed the APSP data). The reason that the second detail is important is: Suppose we have another routine that given an r -truncated DSO \mathcal{D} , constructs $\text{Extend}(\mathcal{D})$ which is a $(3/2)r$ -truncated DSO with a possibly large query time. Then given an 1-truncated DSO $\mathcal{D}^{\text{start}}$, we can construct a (normal) DSO as follows:

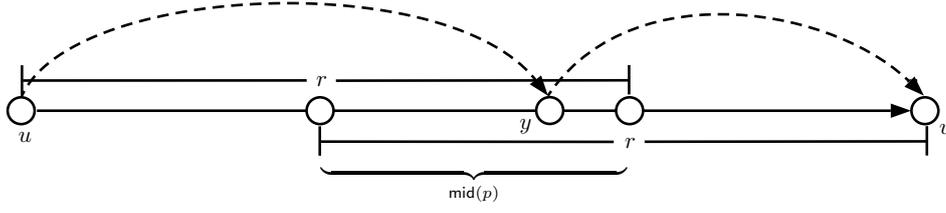
$$\mathcal{D}^{\text{final}} = \underbrace{\text{Fast}(\text{Extend}(\text{Fast}(\text{Extend}(\dots \text{Extend}(\mathcal{D}^{\text{start}})))))}_{O(\log n) \text{ times}}.$$

However, even if the preprocessing algorithm of $\text{Fast}(\mathcal{D})$ invokes the preprocessing algorithm of \mathcal{D} *twice*, the preprocessing algorithm of $\mathcal{D}^{\text{final}}$ would invoke *a polynomial times* the preprocessing algorithm of $\mathcal{D}^{\text{start}}$, which is too many. In contrast, if the preprocessing algorithm of both $\text{Fast}(\mathcal{D})$ and $\text{Extend}(\mathcal{D})$ only invokes the preprocessing algorithm of \mathcal{D} once, then the preprocessing algorithm of $\mathcal{D}^{\text{final}}$ would also invoke the preprocessing algorithm of $\mathcal{D}^{\text{start}}$ only once, which is okay.

2.3 Case II: r is Large

Suppose we have an r -truncated DSO \mathcal{D} , which has preprocessing time P and query time $O(1)$. We show how to construct a $(3/2)r$ -truncated DSO, which we name as $\text{Extend}(\mathcal{D})$, with preprocessing time $P + O(n^2)$ and query time $\tilde{O}(nM/r)$. This is done by the following bridging set argument.

Let \mathcal{P} be the set of paths $uv \diamond f$, for every $u, v \in V$ and $f \in (V \cup E)$ such that $r \leq \|uv \diamond f\| < (3/2)r$. This corresponds to the paths that \mathcal{D} cannot deal with, but $\text{Extend}(\mathcal{D})$ has to output the correct answer. Let $p = uv \diamond f \in \mathcal{P}$, $\text{mid}(p)$ be the set of vertices $y \in p$ such that $\|p[u, y]\| < r$ and $\|p[y, v]\| < r$. (See Figure 1.) For every $y \in \text{mid}(p)$, as $p[u, y] = uy \diamond f$ and $p[y, v] = yv \diamond f$, it follows that \mathcal{D} can find $\|uy \diamond f\|$ and $\|yv \diamond f\|$ correctly. Moreover, $|\text{mid}(p)| \geq r/3M$.



■ **Figure 1** Example of a path $p = uv \diamond f$. If we can find a vertex $y \in \text{mid}(p)$, then we can use \mathcal{D} to compute $\|uy \diamond f\|$ and $\|yv \diamond f\|$, thus to compute the length of p .

Fix a large enough constant C , the preprocessing algorithm of $\text{Extend}(\mathcal{D})$ is as follows: We preprocess \mathcal{D} , and then randomly sample a set H of vertices, where every vertex $v \in V$ is in H with probability $\min\{1, 3CM \ln n/r\}$ independently. We have $|H| = \tilde{O}(nM/r)$ w.h.p.

Fix $u, v \in V$ and $f \in (V \cup E)$, suppose $p = uv \diamond f$ and $r \leq \|p\| < (3/2)r$. Then the probability that H hits $\text{mid}(p)$ (i.e. $H \cap \text{mid}(p) \neq \emptyset$) is at least

$$1 - (1 - 3CM \ln n/r)^{r/3M} \geq 1 - 1/n^C.$$

By a union bound over $O(n^4)$ paths in \mathcal{P} , it follows that w.h.p. H hits $\text{mid}(p)$ for every path $p \in \mathcal{P}$.

The query algorithm for $\text{Extend}(\mathcal{D})$ is as follows: Given a query (u, v, f) , if $\mathcal{D}(u, v, f) < r$, then we output $\mathcal{D}(u, v, f)$; otherwise we output

$$\min \left\{ (3/2)r, \min_{h \in H, \mathcal{D}(u, h, f) < r, \mathcal{D}(h, v, f) < r} \{ \mathcal{D}(u, h, f) + \mathcal{D}(h, v, f) \} \right\}.$$

It is easy to see that $\text{Extend}(\mathcal{D})$ is a correct $(3/2)r$ -truncated DSO, has preprocessing time $P + O(n^2)$ and query time $\tilde{O}(nM/r)$.

2.4 Putting It Together

Let $a \in [0, 1]$ be a constant that we pick later, and $r = n^a$. We first compute the APSP data using Theorem 1.4, which costs $\tilde{O}(n^{(3+\omega)/2} M^{1/2})$ time and will not be the bottleneck. Then we invoke Section 2.1 to build an r -truncated DSO for $r = n^a$, which costs $\tilde{O}(r^2 M \cdot \text{MM}(n, n/r, n))$ time for directed graphs or $\tilde{O}(r \cdot Mn^\omega)$ for undirected graphs. Then for every $1 \leq i \leq \lceil \log_{3/2}(Mn/r) \rceil$, suppose we have an $r(3/2)^{i-1}$ -truncated DSO \mathcal{D}^{i-1} , we can construct $\mathcal{D}^i = \text{Fast}(\text{Extend}(\mathcal{D}^{i-1}))$ which is an $r(3/2)^i$ -truncated DSO. This step costs $\tilde{O}(n^3 M / (r(3/2)^i))$ time. The preprocessing algorithm terminates when $i = i_* = \lceil \log_{3/2}(Mn/r) \rceil = O(\log n)$, and we obtain an $r(3/2)^{i_*}$ -truncated DSO which is a (normal) DSO.

Case 1: the input graph is undirected. The total preprocessing time is

$$\tilde{O}(r \cdot Mn^\omega + n^3 M/r) = \tilde{O}(n^{\max\{\omega+a, 3-a\}} M).$$

When $a = (3 - \omega)/2$, this time complexity is $\tilde{O}(n^{(3+\omega)/2} M) = \tilde{O}(n^{2.6865} M)$.

Therefore, given an undirected graph $G = (V, E)$, there is a DSO with $\tilde{O}(n^{2.6865} M)$ preprocessing time and $O(1)$ query time.

Case 2: the input graph is directed. The total preprocessing time is

$$\tilde{O}(r^2 M \cdot \text{MM}(n, n/r, n) + n^3 M/r) = \tilde{O}(n^{2a+\omega(1-a)} M + n^{3-a} M).$$

(Recall that $\omega(1-a)$ is the exponent of multiplying an $n \times n^{1-a}$ matrix and an $n^{1-a} \times n$ matrix.)

Let $a = 0.276724$, then $1-a = 0.723276$. By convexity of the function $\omega(\cdot)$ [15], we have

$$\omega(1-a) \leq \frac{(a-0.25)\omega(0.7) + (0.3-a)\omega(0.75)}{0.75-0.7}.$$

We substitute $\omega(0.7) \leq 2.154399$ and $\omega(0.75) \leq 2.187543$ [12], and obtain:

$$\omega(1-a) \leq 20 \cdot ((a-0.25) \cdot 2.154399 + (0.3-a) \cdot 2.187543) \leq 2.169829.$$

Therefore, given a directed graph $G = (V, E)$, there is a DSO with

$$\tilde{O}(n^{\max\{2a+\omega(1-a), 3-a\}} M) = \tilde{O}(n^{2.723277} M).$$

preprocessing time and $O(1)$ query time.

3 Proof of Observation 2.1 and Observation 1.1

In this section, we prove Observation 2.1. Note that Observation 1.1 follows from Observation 2.1 by setting $r = +\infty$.

► **Observation 2.1.** *Let r be an integer, $G = (V, E)$ be an input graph, and \mathcal{D} be an arbitrary r -truncated DSO. We can construct $\text{Fast}(\mathcal{D})$, which is an r -truncated DSO with $O(1)$ query time and a preprocessing algorithm as follows.*

- *It first computes the distance matrix of G , and the (incoming and outgoing) shortest path trees rooted at each vertex.*
- *Then it invokes the preprocessing algorithm of \mathcal{D} on the input graph G .*
- *At last, it makes $\tilde{O}(n^2)$ queries to \mathcal{D} , and spends $\tilde{O}(n^2)$ additional time to finish the preprocessing algorithm.*

3.1 The Preprocessing Algorithm

We review and slightly modify the preprocessing algorithm of [3]. For convenience, we denote $\|p\|_r = \min\{\|p\|, r\}$ for any path p and number r .

Assigning priorities. We assign each vertex a *priority*, which is independently sampled from the following distribution: for any positive integer c , each vertex has priority c w.p. $1/2^c$. Denote $c(v)$ the priority of the vertex v . With high probability, all of the following are true:

- The maximum priority is $O(\log n)$.
- For every $c \leq O(\log n)$, there are $\tilde{O}(n/2^c)$ vertices with priority c .
- Let C be a large enough constant. For every shortest path uv with at least $C \cdot 2^c \log n$ edges, there is a vertex on uv whose priority is greater than c .

In the following discussions, we will assume that all of the above assumptions hold.

Fix a pair $u, v \in V$, let s_i be the first vertex in uv with priority $\geq i$, and t_i be the last such vertex. Then we can write the path uv as

$$u \rightsquigarrow s_1 \rightsquigarrow s_2 \rightsquigarrow \dots \rightsquigarrow s_{O(\log n)} \rightsquigarrow t_{O(\log n)} \rightsquigarrow \dots \rightsquigarrow t_1 \rightsquigarrow v.$$

We say that the vertices u, v, s_i, t_i are *key vertices*, and the i -th key vertex is denoted as k_i . Then the path uv can also be written as

$$u = k_0 \rightsquigarrow k_1 \rightsquigarrow \dots \rightsquigarrow k_{O(\log n)} = v.$$

It is important to see that

$$|k_i k_{i+1}| \leq C \cdot 2^{\min\{c(k_i), c(k_{i+1})\}} \log n \quad (1)$$

for every valid i , as otherwise there will be another key vertex between k_i and k_{i+1} .

Data structures for quick location. Suppose we are given a query (u, v, f) , the first thing we should do is to “locate” f , i.e. find the key vertices $k_i, k_{i+1} \in uv$ such that $f \in k_i k_{i+1}$. We will utilize the following data (see also [2]).

- $\text{CL}[u, v, c]$ (for “center left”): the first vertex in uv with priority at least c ;
- $\text{CR}[u, v, c]$ (for “center right”): the last vertex in uv with priority at least c ; and
- $\text{BCP}[u, v]$ (for “biggest center priority”): the maximum priority of any vertex on uv .

It is easy to compute these data in $\tilde{O}(n^2)$ time: for every $u \in V$, we perform a depth-first search on the outgoing shortest path tree $T^{\text{out}}(u)$ to compute $\text{BCP}[u, \cdot]$; for every $u \in V$ and each priority c , we also perform a depth-first search on $T^{\text{out}}(u)$ to compute $\text{CL}[u, \cdot, c]$ and $\text{CR}[u, \cdot, c]$.

In addition, for every $u, v \in V$, we store the key vertices on uv into a hash table of size $O(\log n)$. Given a (vertex) failure f , we can output whether f is among these key vertices on uv in $O(1)$ worst-case time [10].

Data structures for avoiding a failure. We use \mathcal{D} to preprocess the input graph. Then we compute the following data:

- (Data a)** For every $u, v \in V$, and every $1 \leq i \leq \min\{C \cdot 2^{c(u)} \log n, |uv|\}$, let x_i be the i -th vertex in the path uv . (Here u is the 0-th vertex.) We compute and store the value $\|uv \diamond x_i\|_r$. Also, let e_i be the edge from x_{i-1} to x_i , we compute and store $\|uv \diamond e_i\|_r$. Symmetrically, let x_{-i} be the *last* i -th vertex in the path vu (not uv !), and e_{-i} be the edge from x_{-i} to $x_{-(i-1)}$. For every $1 \leq i \leq \min\{C \cdot 2^{c(u)} \log n, |vu|\}$, we compute and store $\|vu \diamond x_{-i}\|_r$ and $\|vu \diamond e_{-i}\|_r$.
- (Data b)** For every $u, v \in V$ and consecutive key vertices $k_i, k_{i+1} \in uv$, let y be the vertex in the portion $k_i \rightsquigarrow k_{i+1}$ that maximizes $\|uv \diamond y\|_r$. We compute and store $\|uv \diamond y\|_r$.
- (Data c)** For every $u, v \in V$ and key vertex $k_i \in uv$, we compute and store $\|uv \diamond k_i\|_r$.

For each priority $c \leq \tilde{O}(1)$, there are $\tilde{O}(n/2^c)$ vertices u whose priority is exactly c . In (Data a), we make $\tilde{O}(n2^c)$ queries for each such u ($\tilde{O}(2^c)$ queries for each $v \in V$). Therefore in total, we make $\tilde{O}(n^2)$ queries in (Data a). We will show in Section 3.3 that we can compute (Data b) using $\tilde{O}(n^2)$ queries to \mathcal{D} and $\tilde{O}(n^2)$ additional time. (Data c) can be computed in $\tilde{O}(n^2)$ queries easily.

3.2 The Query Algorithm

Let (u, v, f) be a query. We first check whether $f \in uv$ in the shortest path trees; if $f \notin uv$, then it is easy to see that $\|uv \diamond f\|_r = \|uv\|_r$.

If f is a vertex failure, we check whether f is a key vertex on uv (that is, $f = k_i$ for some i), using the hash tables. If this is the case, we return $\|uv \diamond f\|_r$ stored in (Data c) immediately.

Otherwise, we start by finding two consecutive key vertices $k_i, k_{i+1} \in uv$ such that $f \in k_i k_{i+1}$. Recall that, if ℓ is the biggest priority of any vertex on uv , then the key vertices on uv are

$$(u =) \text{CL}[u, v, 1] \rightsquigarrow \text{CL}[u, v, 2] \rightsquigarrow \dots \rightsquigarrow \text{CL}[u, v, \ell] \rightsquigarrow \text{CR}[u, v, \ell] \rightsquigarrow \dots \rightsquigarrow \text{CR}[u, v, 2] \rightsquigarrow \text{CR}[u, v, 1](= v).$$

Denote a and b as the “tail” and “head” of f respectively. In particular, if f is a vertex failure then $a = b = f$; if f is an edge failure then it is an edge that goes from a to b . We can find k_i in $O(1)$ time using the following procedure:

- If $\text{BCP}[b, v] = \ell$, then f is in the range $(u, \text{CR}[u, v, \ell])$, so $k_i = \text{CL}[u, v, \text{BCP}[u, a]]$.
- Otherwise, f is in the range $(\text{CR}[u, v, \ell], v)$ and we can see that $k_i = \text{CR}[u, v, \text{BCP}[b, v] + 1]$.

We can find k_{i+1} similarly. By Eq. (1), if $k_i = u$, then $|ub| \leq C \cdot 2^{c(u)} \log n$, and we can look up the value $\|uv \diamond f\|_r$ from (Data a) directly. Similarly, if $k_{i+1} = v$ then we can also look up $\|uv \diamond f\|_r$ from (Data a).

Now we assume that $k_i \neq u$ and $k_{i+1} \neq v$. A crucial observation is that

$$\|uv \diamond f\| = \min\{\|uk_{i+1} \diamond f\| + \|k_{i+1}v\|, \|uk_i\| + \|k_i v \diamond f\|, \|uv \diamond y\|\}, \quad (2)$$

where y is the vertex in $[k_i, k_{i+1}]$ that maximizes $\|uv \diamond y\|$. The proof of Eq. (2) is as follows:

- (i) If $uv \diamond f$ goes through k_i , then $\|uv \diamond f\| = \|uk_i\| + \|k_i v \diamond f\|$.
- (ii) If $uv \diamond f$ goes through k_{i+1} , then $\|uv \diamond f\| = \|uk_{i+1} \diamond f\| + \|k_{i+1}v\|$.
- (iii) If $uv \diamond f$ goes through neither k_i nor k_{i+1} , then it avoids the entire portion of $k_i \rightsquigarrow k_{i+1}$, thus also avoids y . We have $\|uv \diamond f\| \geq \|uv \diamond y\|$. But $\|uv \diamond y\| \geq \|uv \diamond a\|$ by definition of y , and $\|uv \diamond a\| \geq \|uv \diamond f\|$. (Recall that a is the “tail” of f .) Thus $\|uv \diamond f\| = \|uv \diamond y\|$.

It is easy to see that a similar equation holds for r -truncated DSOs:

$$\|uv \diamond f\|_r = \min\{\|uk_{i+1} \diamond f\|_r + \|k_{i+1}v\|, \|uk_i\| + \|k_i v \diamond f\|_r, \|uv \diamond y\|_r, r\}, \quad (3)$$

where y is any vertex in $[k_i, k_{i+1}]$ that maximizes $\|uv \diamond y\|_r$.

Recall that we already know the values $\|uk_i\|$ and $\|k_{i+1}v\|$. To compute $\|uk_{i+1} \diamond f\|_r$, we note that if f is the last j -th vertex or edge in uk_{i+1} , then $j \leq C \cdot 2^{c(k_{i+1})} \log n$. Therefore we can look up the value of $\|uk_{i+1} \diamond f\|_r$ from (Data a). Similarly we can look up $\|k_i v \diamond f\|_r$. Finally, we can look up $\|uv \diamond y\|_r$ from (Data b).

We can see that the query time is $O(1)$.

3.3 Computing (Data b)

We will use the following notation. Let p be a path from u to v which is fixed in context, and a, b be two vertices in p . We will say that $a < b$ if $|p[u, a]| < |p[u, b]|$, i.e. a appears strictly before b on the path p . Similarly, $a > b$, $a \leq b$, $a \geq b$ mean $|p[u, a]| > |p[u, b]|$, $|p[u, a]| \leq |p[u, b]|$, $|p[u, a]| \geq |p[u, b]|$ respectively.

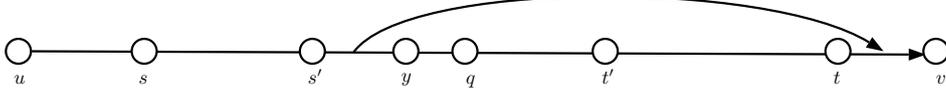
Let $u, v \in V$ and $s < t$ be two vertices on the path uv . Let $y \in [s, t]$ be the vertex in $[s, t]$ which maximizes $\|uv \diamond y\|_r$. We first show that *assuming we have built some oracles*, we can find this vertex y in $O(\log n)$ oracle calls and $O(\log n)$ additional time. The idea is to use a binary search described in [3, Section 6].

► **Lemma 3.1.** *Let r be an integer, $u, v \in V$, and s, t be two vertices on the path uv such that $u < s < t < v$. Suppose we have the following oracles, each with $O(1)$ query time:*

- an oracle that given a vertex $x \in st$, outputs $\|ut \diamond x\|_r$;
- an oracle that given a vertex $x \in st$, outputs $\|sv \diamond x\|_r$;

- an oracle that given an interval $[s', t']$ such that $s \leq s' \leq t' \leq t$, outputs a vertex $x \in [s', t']$ that maximizes the value $\|ut \diamond x\|_r$.

Then we can find a vertex $y \in [s, t]$ which maximizes $\|uv \diamond y\|_r$ in $O(\log n)$ time.



■ **Figure 2** If $sv \diamond y$ does not go through t , then $sv \diamond y$ does not go through the whole interval $[q, t']$.

Proof. For any $y \in [s, t]$, we denote

$$h(y) = \min\{\|ut \diamond y\|_r + \|tv\|, \|us\| + \|sv \diamond y\|_r, r\}.$$

By Eq. (3), we have $\|uv \diamond y\|_r = \min\{h(y), \|uv \diamond y^*\|_r\}$ where y^* is some vertex independent of y . Thus it suffices to find some $y \in [s, t]$ that maximizes $h(y)$.

We use a binary search. Assume that we know the optimal y is in some interval $[s', t']$, where $s \leq s' < t' \leq t$. (Initially we set $s' = s$ and $t' = t$.) If $|s't'| = O(1)$ then we can use brute force to find a vertex $y \in [s', t']$ that maximizes $h(y)$. Otherwise let q be the middle point of $[s', t']$, and we use the third oracle to find a vertex $y \in [s', q]$ that maximizes $\|ut \diamond y\|_r$. There are two cases:

- If $\min\{\|ut \diamond y\|_r + \|tv\|, r\} = h(y)$, then we can restrict our attention to the interval $[q, t']$. This is because for every vertex $x \in [s', q]$,

$$h(x) \leq \min\{\|ut \diamond x\|_r + \|tv\|, r\} \leq \min\{\|ut \diamond y\|_r + \|tv\|, r\} = h(y).$$

- Otherwise, $h(y) = \|us\| + \|sv \diamond y\|_r$, and $sv \diamond y$ does not go through t . Therefore $sv \diamond y$ avoids every vertex in $[q, t']$. (See Figure 2.) For every vertex $x \in [q, t']$,

$$h(x) \leq \|us\| + \|sv \diamond x\|_r \leq \|us\| + \|sv \diamond y\|_r = h(y).$$

It follows that we can restrict our attention to the interval $[s', q]$ now.

Therefore, we can always shrink the length of our candidate interval $[s', t']$ by a half. It follows that we can find the desired vertex y in $O(\log n)$ time. ◀

Now we show how to compute (Data b) in $\tilde{O}(n^2)$ time (assuming that (Data a) is ready). The most crucial ingredient is the following Range Maximum Query (RMQ) structures (used in the third item of Lemma 3.1).

For every $u, v \in V$, consider the following sequence (of length $\ell = \min\{|uv| - 1, C \cdot 2^{c(v)} \log n\}$):

$$(\|uv \diamond x_{-1}\|_r, \|uv \diamond x_{-2}\|_r, \dots, \|uv \diamond x_{-\ell}\|_r),$$

where x_{-i} denotes the last i -th vertex in the path uv (v is the last 0-th). We build an RMQ structure of this sequence, which given a query (s, t) ($1 \leq s \leq t \leq \ell$), outputs a number $i \in [s, t]$ that maximizes $\|uv \diamond x_{-i}\|_r$. After we compute the above sequence, this data structure can be preprocessed in $O(\ell)$ time, and each query costs $O(1)$ time [1].

For every priority $c \leq O(\log n)$, there are $\tilde{O}(n/2^c)$ vertices v of this priority, and for each vertex v we construct n RMQ structures (one for each $u \in V$) on length- $\tilde{O}(2^c)$ sequences. The total size of these RMQ structures is

$$\sum_{c=1}^{O(\log n)} \tilde{O}(n/2^c) \cdot n \cdot \tilde{O}(2^c) = \tilde{O}(n^2).$$

Therefore, these RMQ structures can be preprocessed in $\tilde{O}(n^2)$ time. (Note that every element $\|uv \diamond x_{-i}\|_r$ is already computed in (Data a).)

To compute (Data b), we enumerate u, v, k_i, k_{i+1} where k_i, k_{i+1} are consecutive key vertices in uv . There are $\tilde{O}(n^2)$ possible combinations of (u, v, k_i, k_{i+1}) . As argued in Section 3.2, we know that the following data are already computed in (Data a):

- $\|uk_{i+1} \diamond x\|_r$, for any $x \in k_i k_{i+1}$;
- $\|k_i v \diamond x\|_r$, for any $x \in k_i k_{i+1}$.

We also have the following RMQ oracles constructed above:

- An oracle that given any interval $[s', t']$ on the path uv such that $k_i \leq s' \leq t' \leq k_{i+1}$, finds the vertex $y \in [s', t']$ that maximizes $\|uk_{i+1} \diamond y\|_r$ in $O(1)$ time.

It follows from Lemma 3.1 that we can find a vertex $y \in [k_i, k_{i+1}]$ that maximizes $\|uv \diamond y\|_r$ in $O(\log n)$ time. The total time for computing (Data b) is thus $\tilde{O}(n^2)$.

4 Open Questions

The main open problem after this work is to improve the preprocessing time for DSOs. Can we improve the preprocessing time for directed graphs to $\tilde{O}(n^{2.5286}M)$, matching the current best algorithm for APSP in directed graphs [22]? Or at least can we achieve a preprocessing time of $\tilde{O}(n^{2.5}M)$, under the assumption that $\omega = 2$?

Also, can we extend our technique to deal with negative weights? There are a few candidate definitions of “ r -truncated DSOs” in this case, if we interpret r as the number of edges in the shortest path, instead of the length of the shortest path. For example, we may define a DSO is r -truncated if on a query (u, v, f) , it outputs some value no less than $\|uv \diamond f\|$, and when $|uv \diamond f| \leq r$, it outputs $\|uv \diamond f\|$ exactly. However, it seems that every definition of “ r -truncated DSO” that we tried were not compatible with arguments in Section 3.3.

It is annoying that the shortest paths are usually not unique in a graph with small integer edge weights. For directed graphs, Theorem 1.4 shows that we can compute a set of incoming and outgoing shortest path trees that satisfies Theorem 1.3 in $\tilde{O}(n^{(3+\omega)/2}M^{1/2})$ time. Can we achieve the same requirements in $\tilde{O}(n^{2.5286}M)$ time, matching the current best algorithm for APSP?

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A Breaking Ties for Non-unique Shortest Paths

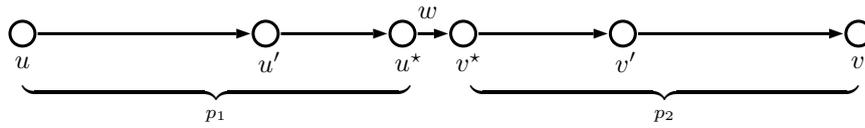
In this section, we prove Theorem 1.3 and 1.4.

Recall that in a subgraph G' of G , we denote $w_{G'}(u, v)$ as the largest (w.r.t. \prec) edge w such that there is a shortest path from u to v whose bottleneck edge (i.e. smallest w.r.t. \prec) is w . Also the path $\rho_{G'}(u, v)$ is defined, inductively from the smallest $\|uv\|$ to the largest $\|uv\|$, as follows: If $u = v$, then $\rho_{G'}(u, v)$ is the empty path; otherwise, suppose $w_{G'}(u, v)$ is an edge from u^* to v^* , then $\rho_{G'}(u, v)$ is the concatenation of $\rho_{G'}(u, u^*)$, $w_{G'}(u, v)$ and $\rho_{G'}(v^*, v)$.

► **Theorem 1.3.** *The following properties of $\rho_{G'}(u, v)$ are true:*

- (Property a) *For every $u', v' \in \rho_{G'}(u, v)$ such that u' appears before v' , the portion of $u' \rightsquigarrow v'$ in $\rho_{G'}(u, v)$ coincides with the path $\rho_{G'}(u', v')$.*
- (Property b) *Let G' be a subgraph of G , suppose $\rho_G(u, v)$ is completely contained in G' , then $\rho_{G'}(u, v) = \rho_G(u, v)$.*

Proof. In this proof we will always use the following notation: Let $w = w_{G'}(u, v)$, and suppose w is an edge from u^* to v^* . Denote $p_1 = \rho_{G'}(u, u^*)$ and $p_2 = \rho_{G'}(v^*, v)$. We will prove Theorem 1.3 inductively, from the smallest $\|uv\|$ to the largest $\|uv\|$. Theorem 1.3 is clearly true when $u = v$.



(Property a): If both $u', v' \in p_1$, then (Property a) is true by induction on p_1 . Similarly, if both $u', v' \in p_2$, then (Property a) is also true by induction on p_2 .

Now suppose $u' \in p_1$ and $v' \in p_2$. Our first observation is that $w_{G'}(u', v') = w_{G'}(u, v) = w$. Actually, since $(\rho_{G'}(u, v))[u', v']$ is a shortest path from u' to v' with bottleneck edge w , we have $w_{G'}(u', v') \succeq w$. But if there is any shortest path p' from u' to v' whose bottleneck edge is $\succ w$, the concatenation of $p_1[u, u']$, p' and $p_2[v', v]$ will be a shortest path from u to v with bottleneck edge $\succ w$, a contradiction. Thus we can only have $w_{G'}(u', v') = w$.

By induction on p_1 , we have $\rho_{G'}(u', u^*) = p_1[u', u^*]$. By induction on p_2 , we have $\rho_{G'}(v^*, v) = p_2[v^*, v]$. Therefore, $\rho_{G'}(u', v')$ is the concatenation of $p_1[u', u^*]$, w , and $p_2[v^*, v]$. This path coincides with the $u' \rightsquigarrow v'$ portion of $\rho_{G'}(u, v)$.

(Property b): it suffices to prove $w_G(u, v) = w_{G'}(u, v)$, and use induction on p_1 and p_2 . Since G' is a subgraph of G , and the distance from u to v is the same in G and G' , we have

$w_G(u, v) \succeq w_{G'}(u, v)$. Since $\rho_G(u, v)$ is completely in G' and the bottleneck edge of $\rho_G(u, v)$ is $w_G(u, v)$, we have $w_{G'}(u, v) \succeq w_G(u, v)$. This completes the proof. \blacktriangleleft

► **Theorem 1.4** ([8]). *Given a subgraph G' , the incoming and outgoing shortest path trees $T^{\text{in}}(u)$ and $T^{\text{out}}(u)$ for each vertex u can be computed in $\tilde{O}(n^{(3+\omega)/2}M^{1/2})$ time.*

Proof Sketch. The first step is to compute $w_{G'}(u, v)$ for every pair of vertices $u, v \in V$. This problem is exactly the same as the *all-pairs bottleneck shortest paths* problem (APBSP; [8, Theorem 4.4]). However, Theorem 4.4 of [8] only claims to work for unweighted graphs, so we verify that it also works for graphs with positive integer weights here.

We start by using [22] to compute $\|\rho_{G'}(u, v)\|$ for every $u, v \in V$. Then, for every $0 \leq i \leq \lceil \log_{3/2}(nM) \rceil$ and $r = (3/2)^i$, assuming we have computed $w_{G'}(u, v)$ for every u, v such that $\|\rho_{G'}(u, v)\| \leq r$, we compute $w_{G'}(u, v)$ for every u, v such that $\|\rho_{G'}(u, v)\| \leq (3/2)r$. This is done as follows. We sample a vertex set H by adding each vertex into H independently w.p. $\Theta(M \log n/r)$, then w.h.p. H hits $\text{mid}(\rho_{G'}(u, v))$ for every u, v such that $r < \|\rho_{G'}(u, v)\| \leq (3/2)r$. (Recall that $\text{mid}(\rho_{G'}(u, v))$ is the “middle 1/3 part” of $\rho_{G'}(u, v)$.) Let D, W be $V \times H$ matrices such that for every $u \in V, h \in H$,

$$D[u, h] = \begin{cases} \|\rho_{G'}(u, h)\| & \|\rho_{G'}(u, h)\| \leq r \\ +\infty & \|\rho_{G'}(u, h)\| > r \end{cases}; \quad W[u, h] = \begin{cases} w_{G'}(u, h) & \|\rho_{G'}(u, h)\| \leq r \\ -\infty & \|\rho_{G'}(u, h)\| > r \end{cases}.$$

Then we compute the Distance-Max-Min product (Definition 4.3 of [8]) of (D, W) and (D^T, W^T) to obtain a $V \times V$ matrix W' such that

$$W'[u, v] = \max_{h: D[u, h] + D[h, v] = \|\rho_{G'}(u, v)\|} \{\min\{W[u, h], W[h, v]\}\}.$$

We can see that $W'[u, v] = w_{G'}(u, v)$ for every $u, v \in V$ such that $r < \|\rho_{G'}(u, v)\| \leq (3/2)r$. This completes the description of the APBSP algorithm.

Now we analyze the time complexity. Let $|H| = n^s$, then $r = \tilde{O}(n^{1-s}M)$, and the finite entries in D are in $\{1, 2, \dots, r\}$. by [8, Theorem 4.3], the Distance-Max-Min product takes

$$O\left(\min\left\{n^{2+s}, n^{1/2}n^{1+s/2+\omega(s)/2}\right\}\right) = \tilde{O}\left(\min\left\{n^{2+s}, n^{(3+\omega(s))/2}M^{1/2}\right\}\right) \leq \tilde{O}\left(n^{(3+\omega)/2}M^{1/2}\right)$$

time. Since we only execute $O(\log n)$ rounds of Distance-Max-Min product, the overall running time of the algorithm is $\tilde{O}(n^{(3+\omega)/2}M^{1/2})$.

Now we construct the outgoing shortest path trees from the table of $w_{G'}(u, v)$ for every $u, v \in V$. It suffices to compute the parents of each node v in the trees $T^{\text{out}}(u)$ (which we denote as $\text{parent}_u(v)$). We compute $\text{parent}_u(v)$ inductively from the smallest $\|uv\|$ to the largest.

Let (u, v) be the pair of vertices we are processing, and assume that for every $u', v' \in V$ such that $\|u'v'\| < \|uv\|$, we have already computed $\text{parent}_{u'}(v')$. Suppose $w_{G'}(u, v)$ is an edge from u^* to v^* . If $v^* = v$, then $\text{parent}_u(v) = u^*$. Otherwise, it is easy to see that $\|\rho_{G'}(v^*, v)\| < \|\rho_{G'}(u, v)\|$ and $\text{parent}_u(v) = \text{parent}_{v^*}(v)$. Thus we can compute every outgoing shortest path tree in $\tilde{O}(n^2)$ time. Similarly, the incoming shortest path trees can also be computed in $\tilde{O}(n^2)$ time. \blacktriangleleft