

Local Algorithms for Sparse Spanning Graphs

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Abstract

We initiate the study of the problem of designing sublinear-time (*local*) algorithms that, given an edge (u, v) in a connected graph $G = (V, E)$, decide whether (u, v) belongs to a sparse spanning graph $G' = (V, E')$ of G . Namely, G' should be connected and $|E'|$ should be upper bounded by $(1 + \epsilon)|V|$ for a given parameter $\epsilon > 0$. To this end the algorithms may query the incidence relation of the graph G , and we seek algorithms whose query complexity and running time (per given edge (u, v)) is as small as possible. Such an algorithm may be randomized but (for a fixed choice of its random coins) its decision on different edges in the graph should be consistent with the same spanning graph G' and independent of the order of queries.

We first show that for general (bounded-degree) graphs, the query complexity of any such algorithm must be $\Omega(\sqrt{|V|})$. This lower bound holds for graphs that have high expansion. We then turn to design and analyze algorithms both for graphs with high expansion (obtaining a result that roughly matches the lower bound) and for graphs that are (strongly) non-expanding (obtaining results in which the complexity does not depend on $|V|$). The complexity of the problem for graphs that do not fall into these two categories is left as an open question.

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1 Introduction

When dealing with large graphs, it is often important to work with a sparse subgraph that maintains essential properties, such as connectivity, bounded diameter and other distance metric properties, of the original input graph. Can one provide fast random access to such a sparsified approximation of the original input graph? In this work, we consider the property of connectivity: Given a connected graph $G = (V, E)$, find a sparse subgraph of G' that spans G . This task can be accomplished by constructing a spanning tree in linear time. However, it can be crucial to *quickly* determine whether a particular edge e belongs to such a subgraph G' , where by “quickly” we mean in time much faster than constructing all of G' . The hope is that by inspecting only some small local neighborhood of e , one can answer in such a way that maintains consistency with the same G' on queries to all edges. We focus on such algorithms, which are of use when we do not need to know the answer for every edge at any single point in time, or if there are several independent processes that want to determine the answer for edges of their choice, possibly in parallel.

If we insist that G' have the *minimum* number of edges sufficient for spanning G , namely, that G' be a spanning tree, then it is easy to see that the task cannot be performed in general without inspecting almost all of G . Interestingly, this is in contrast to the seemingly related problem of estimating the weight of a minimum spanning tree in sublinear-time, which can be performed with complexity that does not depend on $n \stackrel{\text{def}}{=} |V|$ [CRT05] (see further discussion in Subsection 1.3.4). To verify this observe that if G consists of a single path, then the algorithm must answer positively on all edges, while if G consists of a cycle, then the algorithm must answer negatively on one edge. However, the two cases cannot be distinguished without inspecting a linear number of edges. If on the other hand we allow the algorithm some more slackness, and rather than requiring that G' be a tree, require that it be relatively sparse, i.e., contains at most $(1 + \epsilon)n$ edges, then the algorithm may answer positively on all cycle edges, so distinguishing between these two cases is no longer necessary.

We thus consider the above relaxed version of the problem and also allow the algorithm a small failure probability (for a precise formal definition, see Section 2). Our first finding (Theorem 1) is that even when allowing this relaxation, for general (bounded-degree) graphs, the algorithm must inspect $\Omega(\sqrt{n})$ edges in G in order to decide for a given e whether it belongs to the sparse spanning graph G' defined by the algorithm. We then turn to design several algorithms and analyze their performance for various families of graphs. The formal statements of our results can be found in Theorems 2, 3, 4, 5, 6 and 10 as well as Corollaries 8 and 9. Here we provide a high-level description of our algorithms and the types of graphs they give meaningful results for.

1.1 Our results

1.1.1 Expanders

The first algorithm we provide, the Centers' Algorithm (which is discussed further in Subsection 1.2), gives meaningful results for graphs in which the neighborhoods of almost all the vertices in the graph expands in a similar rate. In particular, for graphs with high expansion we get query and running time complexity nearly $O(n^{1/2})$. Since our lower bound applies for graphs with high expansion we obtain that for these graphs, our algorithm is nearly optimal in terms of the complexity in n . More specifically, if the expansion of small sets (of size roughly $O(n^{1/2})$) is $\Omega(d)$, where d is the maximum degree in the graph, then the complexity of the algorithm is $n^{1/2+O(1/\log d)}$. In general, we obtain a sublinear complexity for graphs with expansion (of small sets) that is at least $d^{1/2+1/\log n}$.

1.1.2 Anti-expanders (hyperfinite graphs) and slowly expanding graphs

A graph is ρ -hyperfinite for a function $\rho : \mathbb{R}_+ \rightarrow \mathbb{R}_+$, if its vertices can be partitioned into subsets of size at most $\rho(\epsilon)$ that are connected and such that the number of edges between the subsets is at most ϵn . For the family of hyperfinite graphs we provide an algorithm, the Kruskal-based algorithm, which has success probability 1 and time and query complexity $O(d^{\rho(\epsilon)})$. In particular, the complexity of the algorithm does not depend on n . (where we assume that $\rho(\epsilon)$ is known).

SUBFAMILIES OF HYPERFINITE GRAPHS. For the subfamily of hyperfinite graphs known as graphs with subexponential-growth, we can estimate the diameter of the sets in the partition and hence replace $\rho(\epsilon)$ with such an estimate. This reduces the complexity of the algorithm when the diameter is significantly smaller than $\rho(\epsilon)$, and removes the assumption that $\rho(\epsilon)$ is known. For the subfamily of graphs with an excluded minor (e.g., planar graphs) we can obtain a quasi-polynomial dependence on d and $1/\epsilon$ by using a *partition oracle* for such graphs [LR13a], and the same technique gives polynomial dependence on these parameter for bounded-treewidth graphs (applying [EHNO11]).

GRAPHS WITH SLOW GROWTH RATE. If we do not require that the algorithm work for every ϵ but rather for some fixed constant ϵ , then the Kruskal-based algorithm gives sublinear complexity under a weaker condition than that defining (ρ -)hyperfinite graphs (in which the desired partition should exist for every ϵ). Roughly speaking, the sizes of the neighborhoods of vertices should have bounded growth rate, where the rate may be exponential but the base of the exponent should be bounded (for details, see Theorem 6).

GRAPHS WITH AN EXCLUDED MINOR - THE WEIGHTED CASE. We also provide a local minimum weight spanning graph algorithm, the Borůvka based algorithm, for *weighted* graphs with an excluded fixed minor. The minimum weight spanning graph problem is a generalization of the sparse spanning graph problem for the weighted case. The requirement is that the weight of the graph G' is upper bounded by $(1 + \epsilon)$ times the minimum weight

of a spanning tree. The time and query complexity of the algorithm are quasi-polynomial in $1/\epsilon$, d and W , where W is the maximum weight of an edge. We use ideas from [LR13a], but the algorithm differs from the abovementioned partition-oracle based algorithm for the unweighted case.

1.2 Our algorithms

On a high-level, underlying each of our algorithms is the existence of a (global) partition of the graph vertices where edges within parts are dealt with differently than edges between parts, either explicitly by the algorithm, or in the analysis. The algorithms differ in the way the partitions are defined, where in particular, the number of parts in the partition may be relatively small or relatively large, and the subgraphs they induce may be connected or unconnected. The algorithms also differ in the way the spanning graph edges are chosen, and in particular whether only some of the edges between parts are selected or possibly all. While one of our algorithms works in a manner that is oblivious of the partition (and the partition is used only in the analysis), the other algorithms need to determine in a local manner whether the end points of the given edge belong to the same part or not, as a first step in deciding whether the edge belongs to the sparse spanning graph.

CENTERS' ALGORITHM. This algorithm is based on the following idea. Suppose we can partition the graph vertices into $\sqrt{\epsilon n}$ disjoint parts where each part is connected. If we now take a spanning tree over each part and select one edge between every two parts that have an edge between them, then we obtain a connected subgraph with at most $(1 + \epsilon)n$ edges. The partition is defined based on $\sqrt{\epsilon n}$ special *center* vertices, which are selected uniformly at random. Each vertex is assigned to the closest center (breaking ties according to a fixed order over the centers), unless it is further than some threshold k from all centers, in which case it is a singleton in the partition. This definition ensures the connectivity of each part.

Given an edge (x, y) , the algorithm finds the centers to which x and y are assigned (or determines that they are singletons). If x and y are assigned to the same center, then the algorithm determines whether the edge between them belongs to the BFS-tree rooted at the center. If they belong to different centers, then the algorithm determines whether (x, y) is the single edge allowed to connect the parts corresponding to the two centers (according to a prespecified rule).¹ If one of them is a singleton, then (x, y) is taken to the spanning graph.

From the above description one can see that there is a certain tension between the complexity of the algorithm and the number of edges in the spanning graph G' . On one hand, the threshold k should be such that with high probability (over the choice of the centers) almost all vertices have a center at distance at most k from them. Thus we need a lower bound (of roughly $\sqrt{n/\epsilon}$) on the size of the distance- k neighborhood of (most) vertices. On the other hand, we also need an upper bound on the size of such neighborhoods so that we can efficiently determine which edges are selected between parts. Hence, this algorithm

¹In fact, it may be the case that for two parts that have edges between them, none of the edges is taken, thus making the argument that the subgraph G' is connected a bit more subtle.

works for graphs in which the sizes of the aforementioned local neighborhoods do not vary by too much and its complexity (in terms of the dependence on n) is $\tilde{O}(\sqrt{n})$. In particular this property holds for good expander graphs. We note that the graphs used in our lower bound construction have this property, so for such graphs we get a roughly tight result.

KRUSKAL BASED ALGORITHM. This algorithm is based on the well known algorithm of Kruskal [Kru56] for finding a minimum weight spanning tree in a weighted graph. We use the order over the edges that is defined by the ids of their endpoints as (distinct) “weights”. This ensures that there is a unique “minimum weight” spanning tree. Here the algorithm simply decides whether to include an edge in the spanning graph G' if it does not find evidence in the distance- k neighborhood of the edge that it is the highest ranking (maximum weight) edge on some cycle.

BORŮVKA BASED ALGORITHM. This algorithm is based on the “Binary Borůvka” algorithm [PR08] for finding a minimum-weight spanning tree. Recall that Borůvka’s algorithm begins by first going over each vertex in the graph and adding the lightest edge adjacent to that vertex. Then the algorithm continues joining the formed clusters in a like manner until a tree spanning all vertices is completed. We aim to locally simulate the execution of Borůvka’s algorithm to a point that on one hand all the clusters are relatively small and on the other hand the number of edges outside the clusters is small. The size of the clusters directly affect the complexity of the algorithm and thus our main challenge is in maintaining these clusters small. To this end we use two different techniques. The first technique is to control the growth of the clusters at each iteration by using a certain random orientation on the edges of the graph. This controls the size of the clusters to some extent. In order to deal with clusters that exceeded the required bound (since the the local simulation is recursive even small deviations can have large impact on the complexity), after each iteration we separate large clusters into smaller ones (here we use the fact that the graph excludes a fixed minor in order to obtain a small separator to each cluster).

1.3 Related work

1.3.1 Local algorithms for other graph problems

The model of *local computation algorithms* as used in this work, was defined by Rubinfeld et al. [RTVX11] (see also [ARVX12]). Such algorithms for maximal independent set, hypergraph coloring, k -CNF and maximum matching are given in [RTVX11, ARVX12, MRVX12, MV13]. This model generalizes other models that have been studied in various contexts, including locally decodable codes (e.g., [MSV99]), local decompression [DLRR13], and local filters/reconstructors [ACCL08, SS10, Bra08, KPS08, JR11, CS06]. Local computation algorithms that give approximate solutions for various optimization problems on graphs, including vertex cover, maximal matching, and other packing and covering problems, can also be derived from sublinear time algorithms for parameter estimation [PR07, MR09, NO08, HKNO09, YYI09].

Campagna et al. [CGR13] provide a local reconstructor for connectivity. Namely, given a graph which is almost connected, their reconstructor provides oracle access to the adjacency matrix of a connected graph which is close to the original graph. We emphasize that our model is different from theirs, in that they allow the addition of new edges to the graph, whereas our algorithms must provide spanning graphs whose edges are present in the original input graph.

1.3.2 Distributed algorithms

The name *local algorithms* is also used in the distributed context [MNS95, NS95, Lin92]. In the distributed *message passing* model, the graph G is both the input to the algorithm and the underlying communication network. Namely, each vertex is assigned a processor, and a processor can send messages to processors that reside on the neighbors of its vertex. The algorithm works in rounds where in each round, each vertex can send messages to all its neighbors. When the algorithm terminates, each processor knows *its part of the answer*. For example, if the task is to compute a (small) vertex cover, then each vertex knows whether or not it is in the vertex cover. If the task is to compute a (close to) minimum weight spanning tree, then each process knows which of the edges incident to its vertex belong to the spanning tree.

As observed by Parnas and Ron [PR07], local distributed algorithms can be used to obtain local computation algorithms as defined in this work, by simply emulating the distributed algorithm on a sufficiently large subgraph of the graph G . However, while the main complexity measure in the distributed setting is the number of rounds (where it is usually assumed that each message is of length $O(\log n)$), our main complexity measure is the number of queries performed on the graph G . By this standard reduction, the bound on the number of queries (and hence running time) depends on the size of the queried subgraph and may grow exponentially with the number of rounds. Therefore, this reduction gives meaningful results only when the number of rounds is significantly smaller than the diameter of the graph.

The problem of computing a minimum weight spanning tree in this model is a central one. Kutten et al. [KP98] provided an algorithm that works in $O(\sqrt{n} \log^* n + D)$ rounds, where D denotes the diameter of the graph. Their result is nearly optimal in terms of the complexity in n , as shown by Peleg et al. [PR00] who provided a lower bound of $\Omega(\sqrt{n}/\log n)$ rounds (when the length of the messages must be bounded).

Another problem studied in the distributed setting that is related to the one studied in this paper, is finding a sparse spanner. The requirement for spanners is much stronger since the distortion of the distance should be as minimal as possible. Thus, to achieve this property, it is usually the case that the number of edges of the spanner is super-linear in n . Pettie [Pet10] was the first to provide a distributed algorithm for finding a low distortion spanner with $O(n)$ edges without requiring messages of unbounded length or $O(D)$ rounds. The number of rounds of his algorithm is $\log^{1+o(1)} n$. Hence, the standard reduction of [PR07] yields a local algorithm with a trivial linear bound on the query complexity.

1.3.3 Local cluster algorithms

Local algorithms for graph theoretic problems have also been given for PageRank computations on the web graph [JW03, Ber06, SBC⁺06, ACL06, ABC⁺08]. Local graph partitioning algorithms have been presented in [ST04, ACL06, AP09, ZLM13, OZ13], which find subsets of vertices whose internal connections are significantly richer than their external connections in time that depends on the size of the cluster that they output. Even when the size of the cluster is guaranteed to be small, it is not obvious how to use these algorithms in the local computation setting where the cluster decompositions must be consistent among queries to all vertices.

1.3.4 Other related sublinear-time approximation algorithms for graphs

The problem of estimating the weight of a minimum weight spanning tree in sublinear time was considered by Chazelle, Rubinfeld and Trevisan [CRT05]. They describe an algorithm whose running time depends on the approximation parameter, the average degree and the range of the weights, but does not directly depend on the number of nodes. A question that has been open since that time, even before local computation algorithms were formally defined, is whether it is possible to quickly determine which edges are in the minimum spanning tree. Our lower bound for spanning trees applies to this question.

Related minimum weight spanning-tree problems, when the graph resides in a metric space, were studied in [CS09, CEF⁺05]. Other papers that deal with sublinear-time approximation algorithms for graph parameters include [PR07, Fei06, GR08, MR09, NO08, GRY11, HKNO09, YYI09, EHNO11].

2 Preliminaries

The graphs we consider have a known degree bound d , and we assume we have query access to their incidence-lists representation. Namely, for any vertex v and index $1 \leq i \leq d$ it is possible to obtain the i^{th} neighbor of v by performing a query to the graph (if v has less than i neighbors, then a special symbol is returned).² If the graph is edge-weighted, then the weight of the edge is returned as well. The number of vertices in the graph is n and we assume that each vertex v has an id, $id(v)$, where there is a full order over the ids.

Definition 1 (Distances and Neighborhoods) *Let $G = (V, E)$ be a graph.*

1. *We denote the distance between two vertices u and v in G by $d_G(u, v)$.*
2. *For vertex $v \in V$ and an integer k , let $\Gamma_k(v, G)$ denote the set of vertices at distance at most k from v and let $C_k(v, G)$ denote the subgraph of G induced by $\Gamma_k(v, G)$.*

²Graphs are allowed to have self-loops and multiple edges, but for our problem we may assume that there are no self-loops and multiple-edges (since the answer on a self-loop can always be negative, and the same is true for all but at most one among a set of parallel edges).

3. Let $n_k(G) \stackrel{\text{def}}{=} \max_{v \in V} |\Gamma_k(v, G)|$.

When the graph G is clear from the context, we shall use the shorthand $d(u, v)$, $\Gamma_k(v)$ and $C_k(v)$ for $d_G(u, v)$, $\Gamma_k(v, G)$ and $C_k(v, G)$, respectively.

Definition 2 (Local Algorithms for sparse spanning graphs) *An algorithm \mathcal{A} is a local sparse spanning graph algorithm if, given parameters $n \geq 1$, $\epsilon \geq 0$ and $0 \leq \delta < 1$ and given query access to the incidence-lists representation of a graph $G = (V, E)$, the algorithm \mathcal{A} provides query access to a subgraph of G , $G' = (V, E')$ such that the following hold:*

1. G' is connected with probability 1.
2. $|E'| \leq (1 + \epsilon) \cdot n$ with probability at least $1 - \delta$, where the probability is taken over the internal coin flips of \mathcal{A} .
3. E' is determined by G and the internal randomness of the oracle.

Namely, on input $(u, v) \in E$, \mathcal{A} returns whether $(u, v) \in E'$ and for any sequence of queries, \mathcal{A} answers consistently with the same G' .

An algorithm \mathcal{A} is a local sparse spanning graph algorithm with respect to a class of graphs \mathcal{C} if the above conditions hold, provided that the input graph G belongs to \mathcal{C} .

We are interested in local algorithms that have small query complexity, namely, that perform few queries to the graph (for each edge they are queried on) and whose running time (per queried edge) is small as well. As for the question of randomness and the implied space complexity of the algorithms, we assume we have a source of (unbounded) public randomness. Under this assumption, our algorithms do not keep a state and a global space is not required. However, if unbounded public randomness is not available, then we note that for our algorithms this is not an issue: One of our algorithms (see Section 5) is actually deterministic, and for the others, the total number of random bits that is actually required (over all possible queries) is upper bounded by the running time of the algorithm, up to a multiplicative factor of $O(\log n)$. In what follows we sometimes describe a global algorithm first, i.e., an algorithm that reads the entire graph and decides the subgraph G' . After that we describe how to locally emulate the global algorithm. Namely on query $e \in E$, we emulate the global algorithm decision on e while performing only a sublinear number of queries.

3 A Lower Bound for General Bounded-Degree Graphs

In this section we prove the following theorem.

Theorem 1 *Any local sparse spanning graph algorithm has query complexity $\Omega(\sqrt{n})$. This result holds for graphs with a constant degree bound d and for constant $0 \leq \epsilon \leq 2d/3$ and $0 \leq \delta < 1/3$.*

Let V be a set of vertices and let v_0 and v_1 be a pair of distinct vertices in V . In order to prove the lower bound we construct two families of random d -regular graphs over V , $\mathcal{F}_{(v_0, v_1)}^+$ and $\mathcal{F}_{(v_0, v_1)}^-$. $\mathcal{F}_{(v_0, v_1)}^+$ is the family of d -regular graphs, $G = (V, E)$, for which $(v_0, v_1) \in E$. $\mathcal{F}_{(v_0, v_1)}^-$ is the family of d -regular for which $(v_0, v_1) \in E$ and the removal of (v_0, v_1) leaves the graph with two connected components, each of size ³ $n/2$. We prove that given (v_0, v_1) , any algorithm that performs at most \sqrt{n}/c queries for some sufficiently large constant $c > 1$ cannot distinguish the case in which the graph is drawn uniformly at random from $\mathcal{F}_{(v_0, v_1)}^+$ from the case in which the graph is drawn uniformly at random from $\mathcal{F}_{(v_0, v_1)}^-$. Essentially, if the number of queries is at most \sqrt{n}/c , then with high constant probability, each new query to the graph returns a new random vertex in both families. By “new vertex” we mean a vertex that neither appeared in the query history nor in the answers history. Since the algorithm must answer consistently with a connected graph G' , for every graph in the support of $\mathcal{F}_{(v_0, v_1)}^-$ it must answer with probability 1 positively on the query (v_0, v_1) . But since the distributions on query-answer histories in both cases are very close statistically, this can be shown to imply that there exist graphs for which the algorithm answers positively on a large fraction of the edges.

Proof of Theorem 1: We consider a pair of oracles that generate a random graph from $\mathcal{F}_{(v_0, v_1)}^+$ and $\mathcal{F}_{(v_0, v_1)}^-$, respectively, on the fly. On query (w, i) the oracle returns (v, j) where v is i -th neighbor of w , denoted $N(w, i)$, and w is the j -th neighbor of v . If w does not have an i -th neighbor then the oracle returns \emptyset . Both oracles construct the graph on demand, i.e. on on query (w, i) , if $N(w, i)$ was not determined in a previous step then the oracle determines $N(w, i)$.

FIRST ORACLE Recall that a random d -regular graph can be generated as follows (the approach was invented by [BC78] and independently by [Bol79]). Define a matrix $M = [n] \times [d]$ where the i -th row corresponds to the i -th vertex. Partition the cells of M into $nd/2$ pairs, namely, find a random perfect matching on the cells of M . Now let (w, i) and (v, j) be a pair of matched cells, then in the corresponding graph it holds that $N(w, i) = v$ and $N(v, j) = w$. Thus, to obtain a random d -regular graph from the family of d -regular graphs $G = (V, E)$, for which $(v_0, v_1) \in E$ proceed as follows.

1. Match a random cell in v_0 's row to a random cell with v_1 's row.
2. Find a random perfect matching of the remaining cells.

³Although a graph that is drawn uniformly from $\mathcal{F}_{(v_0, v_1)}^+$ (or $\mathcal{F}_{(v_0, v_1)}^-$) might be disconnected, this event happens with negligible probability [Bol01]. Hence, the proof of the lower bound remains valid even if we consider $\mathcal{F}_{(v_0, v_1)}^+ \cap \mathcal{C}$ and $\mathcal{F}_{(v_0, v_1)}^- \cap \mathcal{C}$ where \mathcal{C} is the family of connected graphs.

Our first oracle, $\mathcal{O}_{(v_0, v_1)}^+$, determines the perfect matching on demand as we describe next. The oracle keeps a matrix, M , with n rows which correspond to the n vertices and d columns, which correspond to d neighbors. At the initialization step:

1. All the cells are initialized to \emptyset .
2. A single edge is determined: $M(v_0, t_0) := (v_1, t_1)$ and $M(v_1, t_1) := (v_0, t_0)$ where t_0 and t_1 are chosen uniformly and independently over $[d]$.

On query (w, i) , the oracle $\mathcal{O}_{(v_0, v_1)}^+$, proceeds as follows:

1. Checks if $M(w, i)$ was determined in a previous step, i.e. if $M(w, i) \neq \emptyset$, if so it returns the matched cell (u, j) .
2. Otherwise, it picks uniformly one of the empty cell in the matrix, (u, j) .
3. Sets $M(w, i) := (u, j)$, $M(u, j) := (w, i)$. Namely, adds the edge (u, w) to the graph.
4. Returns (u, j) .

Without loss of generality assume that $n \cdot d$ is even ⁴. Thus, if the oracle is queried on all the entries of the matrix then the resulting graph is a random d -regular graph that contain the edge (v_0, v_1) (where self-loops and parallel edges are allowed). Notice that the resulting graph is drawn uniformly (and independently of the order of the queries) from the family of graphs $\mathcal{F}_{(v_0, v_1)}^+$.

SECOND ORACLE Our second oracle, $\mathcal{O}_{(v_0, v_1)}^-$ generates a random graph from the family of graphs $\mathcal{F}_{(v_0, v_1)}^-$. The oracle $\mathcal{O}_{(v_0, v_1)}^-$ keeps a pair of matrices M_0, M_1 . The matrix, M_0 , has $\lfloor n/2 \rfloor$ rows and d columns and the matrix, M_1 , has $\lceil n/2 \rceil$ rows and d columns ⁵. At the initialization step:

1. All the cells of both matrices are initialized to \emptyset .
2. The rows of the matrices are not allocated to any vertex.
3. A random row in M_b, i_b , is allocated to v_b for each $b \in \{0, 1\}$.
4. A single edge is determined: $M_0(i_0, t_0) := M_1(i_1, t_1)$ and $M_1(i_1, t_1) := M_0(i_0, t_0)$ where t_0 and t_1 are chosen uniformly and independently over $[d]$.

For a vertex v , let $r(v)$ denote the index of the row that is allocated to v and the corresponding matrix by M_v . Given a row j and a matrix M , let $v(M, j)$ denote the vertex that the j -th row in M is allocated to. On query (w, i) , $\mathcal{O}_{(v_0, v_1)}^-$ proceeds as follows:

⁴In case that $d \cdot n$ is odd we add an extra entry to the matrix, $(0, 0)$, so that $M(w, i) = (0, 0)$ means that w does not have an i -th neighbor.

⁵Here too, if the number of entries in a matrix is odd we add an entry $(0, 0)$.

1. If a row was not allocated to w in previous steps then a random row is picked uniformly (from the set of rows that are free) and is allocated to w .
2. If the the i -th neighbor of w was determined previously, i.e., the cell $M_w(r(w), i)$ was matched in previous steps to another cell $M_w(\ell, j)$, then return $(v(M_w, \ell), j)$.
3. Otherwise, select randomly and uniformly an empty cell in M_w , (ℓ, j) . If the ℓ -th row of M_w is free, then allocate this row to a vertex which is picked uniformly from the set of vertices without an allocated row.
4. Add an edge between (w, i) and $(v(M_w, \ell), j)$, i.e., set $M_w(r(w), i) := M_w(\ell, j)$ and $M_w(\ell, j) = M_w(r(w), i)$. Return $(v(M_w, \ell), j)$.

Let \mathcal{A} be an algorithm that interacts with $\mathcal{O}_{(v_0, v_1)}^+$ and let $\pi_r^+ = (q_1^+, a_1^+, \dots, q_r^+, a_r^+)$ be the random variable describing the communication between \mathcal{A} and $\mathcal{O}_{(v_0, v_1)}^+$. Namely, π_r^+ is a list of r queries of \mathcal{A} and r corresponding answers of $\mathcal{O}_{(v_0, v_1)}^+$. Similarly, define π_r^- for $\mathcal{O}_{(v_0, v_1)}^-$. We claim that for $r = c\sqrt{n}$, the distribution of π_r^+ is statistically close to the distribution of π_r^- . Recall that every query q_i is a pair $(q_{i,1}, q_{i,2})$ where the first entry denotes a name of a vertex and the second entry denotes an index in $[d]$. Similarly, every answer a_i is a pair $(a_{i,1}, a_{i,2})$. Let A^+ be the event that for every $i \in [r]$ it holds that, $a_{i,1}^+ \neq a_{j,1}^+$ for every $i \neq j \in [r]$ and $a_{i,1}^+ \neq q_{j,1}^+$ for every $j \in [r]$. Similarly, define A^- . Without loss of generality we assume that \mathcal{A} does not make queries that does not reveal any new information on the graph. Hence, in words, A^+ (respectively, A^-) is the event that whenever $\mathcal{O}_{(v_0, v_1)}^+$ (respectively $\mathcal{O}_{(v_0, v_1)}^-$) selects an empty (un-matched) cell, the corresponding row is empty as well. It is easy to verify that the distribution of π_r^+ conditioned on A^+ is identical to the distribution of π_r^- conditioned on A^- . Hence, the statistical distance between the distribution of π_r^+ and the distribution of π_r^- is bounded by $2|\Pr(A^+) - \Pr(A^-)|$. Since both $\Pr(A^+)$ and $\Pr(A^-)$ are upper bounded by $2r \cdot \frac{2r}{n/2} = 8r^2/n = 8c^2$, we obtain that the statistical distance is bounded by $16c^2 = 16/49 < 1/3$ for $c = 1/7$.

From this the lower bound is implied as follows. Assume towards contradiction that there exists a local spanning graph algorithm, \mathcal{A} , with query complexity $c\sqrt{n}$. When \mathcal{A} interacts with $\mathcal{O}_{(v_0, v_1)}^-$ it must answer yes with probability 1 (over the random coins of \mathcal{A} and $\mathcal{O}_{(v_0, v_1)}^-$) when queried on (v_0, v_1) . Since $\mathcal{O}_{(v_0, v_1)}^-$ and $\mathcal{O}_{(v_0, v_1)}^+$ are statistically close, \mathcal{A} that interacts with $\mathcal{O}_{(v_0, v_1)}^+$ must answer yes with probability at least $2/3$ (over the random coins of $\mathcal{O}_{(v_0, v_1)}^-$ and for every setting of random coins of \mathcal{A}) when queried on (v_0, v_1) . This implies that for every setting of random coins of \mathcal{A} , \mathcal{A} answers yes on at least $2/3$ of the d -regular graphs that contain (v_0, v_1) . By an averaging argument it follows that there exists a d -regular graph, G , such that \mathcal{A} answers yes on at least $2/3$ of the edges of G . A contradiction. ■

4 Graphs with High Expansion

In this section we describe an algorithm that gives meaningful results for graphs in which, roughly speaking, the local neighborhood of almost all vertices expands in a similar rate. In particular this includes graphs with high expansion. In fact we only require that the graph expands quickly for small sets: A graph G is an (s, α) -vertex expander if for all sets S of size at most s , $N(S)$ is of size at least $\alpha|S|$, where $N(S)$ denotes the set of vertices adjacent to vertices in S that are not in S . Define $h_s(G)$ to be the maximum α such that G is an (s, α) -vertex expander. We shall prove the following theorem.

Theorem 2 *Given a graph $G = (V, E)$ with degree bound d , there is a local sparse spanning graph algorithm with query complexity and running time $(d \cdot s)^{\log_{h_s(G)} d}$ where $s = s(n, \epsilon, \delta) \stackrel{\text{def}}{=} \sqrt{2n/\epsilon} \cdot \log(n/\delta)$.*

By Theorem 2, for bounded degree graphs with high expansion we get query and running time complexity nearly $O(n^{1/2})$. In particular, if $h_s(G) = \Omega(d)$ for $s = s(n, \epsilon, \delta)$ then the complexity is $n^{1/2+O(1/\log d)}$. In fact, even for $h_s(G) \geq d^{1/2+1/\log n}$ the complexity is $o(n)$. Recall that in the construction of our lower bound of $\Omega(n^{1/2})$ we construct a pair of families of d -regular random graphs. In both families, the expansion (of small sets) is $\Omega(d)$, implying that the complexity of our algorithm is almost tight.

Our algorithm, the local Centers' Algorithm (which appears as Algorithm 1), is based on a global algorithm which is presented in Subsection 4.1. The local Centers' Algorithm appears in Subsection 4.2 and it is analyzed in the proof of Theorem 3.

4.1 The Global Algorithm

For a given parameter k the global algorithm first defines a global partition of (part or all of) the graph vertices in the following randomized manner.

1. Select $\ell = \sqrt{\epsilon n/2}$ centers uniformly and independently at random from V , and denote them v_1, \dots, v_ℓ .
2. Initially, all vertices are *unassigned*.
3. For $i = 0, \dots, k$, for $j = 1, \dots, \ell$:
 Let L_j^i denote the vertices in the i^{th} level of the BFS tree of v_j (where $L_j^0 = \{v_j\}$).
 Assign to v_j all vertices in L_j^i that were not yet assigned to any other $v_{j'}$.

Let $S(v_j)$ denote the set of vertices that are assigned to the center v_j . By the above construction, the subgraph induced by $S(v_j)$ is connected.

The subgraph $G' = (V, E')$ is defined as follows.

1. For each center v , let $E'(v)$ denote the edges of a BFS-tree that spans the subgraph induced by $S(v)$ (where the BFS-tree is determined by the order over the ids of the vertices in $S(v)$). For each center v , put in E' all edges in $E'(v)$.
2. For each vertex w that does not belong to any $S(v)$ for a center v , put in E' all edges incident to w .
3. For each pair of centers u and v , let $P(u, v)$ be the shortest path between u and v that has minimum lexicographic order among all shortest paths (as determined by the ids of the vertices on the path). If all vertices on this path belong either to $S(u)$ or to $S(v)$, then add to E' the single edge $(x, y) \in P(u, v)$ such that $x \in S(u)$ and $y \in S(v)$, where we denote this edge by $e(u, v)$.

In what follows we shall prove that G' is connected and that for k that is sufficiently large, G' is sparse with high probability as well. We begin by proving the latter claim. To this end we define a parameter which determines the minimum distance needed for most vertices to see roughly \sqrt{n} vertices. More formally, define $k_{\epsilon, \delta}^C(G)$ to be the minimum distance k ensuring that all but an $\epsilon/(2d)$ -fraction of the vertices have at least $s(n, \epsilon, \delta)$ vertices in their k -neighborhood. That is,

$$k_{\epsilon, \delta}^C(G) \stackrel{\text{def}}{=} \min_k \{|\{v : \Gamma_k(v) \geq s(n, \epsilon, \delta)\}| \geq (1 - \epsilon/(2d)) |V|\} . \quad (1)$$

We next establish that for $k \geq k_{\epsilon, \delta}^C(G)$ it holds that $|E'| \leq (1 + \epsilon)n$ with probability at least $1 - \delta$, over the random choice of centers. Since for $j = 1, \dots, \ell$ the sets $E'(v_j)$ are disjoint, we have that $|\bigcup_{j=1}^{\ell} E'(v_j)| < n$. Since there is at most one edge $e(u, v)$ added to E' for each pair of centers u, v and the number of centers is $\ell = \sqrt{\epsilon n/2}$, the total number of these edges in E' is at most $\epsilon n/2$. Finally, Let $T \subseteq V$ denote the subset of the vertices, v , such that $|\Gamma_k(v)| \geq s(n, \epsilon, \delta)$. Since the centers are selected uniformly, independently at random, for each $w \in T$ the probability that no vertex in $\Gamma_k(w)$ is selected to be a center is at most $(1 - \log(n/\delta)/\sqrt{\epsilon n/2})^{\sqrt{\epsilon n/2}} < \delta/n$. By taking a union bound over all vertices in T , with probability at least $1 - \delta$, every $w \in T$ is assigned to some center v . Since the number of vertices in $V \setminus T$ is at most $(\epsilon d/2)n$ and each contributes at most d edges to E' , we get the desired upper bound on $|E'|$.

It remains to establish that G' is connected. To this end it suffices to prove that there is a path in G' between every pair of centers u and v . This suffices because for each vertex w that is assigned to some center v , there is a path between w and v (in the BFS-tree of v), and for each vertex w that is not assigned to any center, all edges incident to w belong to E' . The proof proceeds by induction on $d(u, v)$ and the sum of the ids of u and v as follows. For the base case consider a pair of centers u and v for which $d(u, v) = 1$. In this case, the shortest path $P(u, v)$ consists of a single edge (u, v) where $u \in S(u)$ and $v \in S(v)$, implying that $(u, v) \in E'$. For the induction step, consider a pair of centers u and v for which $d(u, v) > 1$, and assume by induction that the claim holds for every pair of centers (u', v') such that either $d(u', v') < d(u, v)$ or $d(u', v') = d(u, v)$ and $id(u') + id(v') < id(u) + id(v)$.

Similarly to base case, if the set of vertices in $P(u, v)$ is contained entirely in $S(u) \cup S(v)$, then u and v are connected by construction. Namely, $P(u, v) = (u, x_1, \dots, x_t, y_s, \dots, y_1, v)$ where $x_1, \dots, x_t \in S(u)$ and $y_1, \dots, y_s \in v$. The edge (x_t, y_s) was added to E' and there are paths in the BFS-trees of u and v between u and x_t and between v and y_s , respectively. Otherwise, we consider two cases.

1. There exists a vertex x in $P(u, v)$, and a center, y , such that $x \in S(y)$. Note that this must be the case when $d(u, v) \leq 2k + 1$. This implies that either $d(x, y) < d(x, v)$ or that $d(x, y) = d(x, v)$ and $id(y) < id(v)$. Hence, either

$$d(u, y) \leq d(u, x) + d(x, y) < d(u, x) + d(x, v) = d(u, v)$$

or $d(u, y) = d(u, v)$ and $id(u) + id(y) < id(u) + id(v)$. In either case we can apply the induction hypothesis to obtain that u and y are connected. A symmetric argument gives us that v and y are connected.

2. Otherwise, all the vertices on the path $P(u, v)$ that do not belong to $S(u) \cup S(v)$ are vertices that are not assigned to any center. Since E' contains all edges incident to such vertices, u and v are connected in this case as well.

4.2 The Local Algorithm

Theorem 3 *Algorithm 1, when run with $k \geq k_{\epsilon, \delta}^C(G)$, is a local sparse spanning graph algorithm. The query complexity and running time of the algorithm are $O(d \cdot n_k(G))$.*

Proof: We prove the theorem by showing that Algorithm 1 is a local emulation of the global algorithm that appears in Subsection 4.1. Given x and y , by performing a BFS to depth k from each of the two vertices, Algorithm 1 either finds the centers u and v that x and y are (respectively) assigned to (by the global algorithm, for the same selection of centers), or for at least one of them it finds no center in the distance k neighborhood. In the latter case, the edge (x, y) belongs to E' , and Algorithm 1 returns a positive answer, as required. In the former case, there are two subcases.

1. If x and y are assigned to the same center, that is, $u = v$, then Algorithm 1 checks whether the edge (x, y) is an edge in the BFS-tree of u (i.e., $(x, y) \in E'(u)$). If x and y are on the same level of the tree (i.e., are at the same distance from u), then Algorithm 1 returns a negative answer, as required. If y is one level further than x , then Algorithm 1 checks whether y has another neighbor w that is also assigned to u , is on the same level as x and has a smaller id than x . Namely, a neighbor of y that is on a shortest path between y and u and has a smaller id than x . If this is the case, then the edge (x, y) does not belong to the tree (but rather the edge (w, y)) so that the algorithm returns a negative answer. If no such neighbor of y exists, then the algorithm returns a positive answer (as required).

Algorithm 1 (Centers' Algorithm)

For a random choice of $\ell = \sqrt{\epsilon n/2}$ centers, v_1, \dots, v_ℓ in V (which is fixed for all queries), and for a given parameter k , on query (x, y) :

1. Perform a BFS to depth k in G from x and from y .
 2. If either $\Gamma_k(x) \cap \{v_1, \dots, v_\ell\} = \emptyset$ or $\Gamma_k(y) \cap \{v_1, \dots, v_\ell\} = \emptyset$, then return YES.
 3. Otherwise, let u be the center closest to x and let v be the center closest to y (if there is more than one such center, break ties according to the order v_1, \dots, v_ℓ).
 4. If $u = v$ then do the following:
 - If $d(x, u) = d(y, u)$, then return NO.
 - If $d(y, u) = d(x, u) + 1$, then consider all neighbors of y , w , on a shortest path between y and u . If there exists such neighbor w for which $id(w) < id(x)$, then return NO, otherwise, return YES.
 5. If $u \neq v$, then perform a BFS of depth k from both of the centers, u and v . Find the shortest path between u and v that has the smallest lexicographical order, and denote it by $P(u, v)$. Return YES if both $x \in P(u, v)$ and $y \in P(u, v)$. Otherwise, return NO.
-

2. If x and y are assigned to different centers, that is, $u \neq v$, then Algorithm 1 determines whether $(x, y) = e(u, v)$ exactly as defined in the global algorithm: The algorithm finds $P(u, v)$ and returns a positive answer if and only if (x, y) belongs to $P(u, v)$. Notice that from the fact that $x \in S(u)$ and $y \in S(v)$ and the fact that (x, y) belongs to $P(u, v)$ it follows that all the vertices on $P(u, v)$ belong to either $S(u)$ or $S(v)$. This is implied from the fact that for every center u and a vertex which is assigned to u , it holds that every vertex on a shortest path between u and w is also assigned to u .

Finally, the bound on the query complexity and running time of Algorithm 1 follows directly by inspection of the algorithm. ■

4.3 The Parameter k

Recall that Algorithm 1 is given a parameter k that determines the depth of the BFS that the algorithm performs. By Theorem 3 it suffices to require that $k \geq k_{\epsilon, \delta}^C(G)$ in order to ensure that the spanning graph obtained by the algorithm is sparse. For the case that k is not given in advance we describe next how to compute k such that with probability at least $1 - \delta$ it holds that

$$k_{\epsilon, \delta}^C(G) \leq k \leq k'_{\epsilon, \delta}(G), \quad (2)$$

where $k'_{\epsilon,\delta}(G) = \min_k \{|\{v : \Gamma_k(v) \geq s(n, \epsilon, \delta)\}| \geq (1 - \frac{\epsilon d}{4}) |V|\}$. Select uniformly at random $s = \Theta(1/\epsilon^2 \log(1/\delta))$ vertices from V . Let v_1, \dots, v_s denote the selected vertices. For each vertex in the sample v_i , let $k_i = \min_k \{\Gamma_k(v_i) \geq s(n, \epsilon, \delta)\}$. Assume without loss of generality that $k_1 \leq \dots \leq k_s$ and set $k = k_{\lceil 1 - \frac{3\epsilon}{8d} \rceil}$. By Chernoff's inequality we obtain that with probability greater than $1 - \delta$ Equation (2) holds.

We are now ready to prove Theorem 2.

Proof of Theorem 2: Assume without loss of generality that $k'_{\epsilon,\delta}(G)$ is unknown and we run Algorithm 1 with k such that $k_{\epsilon,\delta}(G) \leq k \leq k'_{\epsilon,\delta}(G)$. From the fact that $n_k(G) \geq \min\{h_s(G)^k, s\}$ we obtain that $k'_{\epsilon,\delta}(G) \leq \frac{\log s}{\log h_s(G)}$ for $s = s(n, \epsilon, \delta)$. On the other hand, since the degree is bounded by d , it holds that $n_k(G) \leq 1 + d^k$. Hence, by Theorem 3 we obtain that the query complexity is bounded by $d \cdot (1 + d^{\frac{\log s}{\log h_s(G)}})$, as desired. ■

5 Hyperfinite Graphs

In this section we provide an algorithm that is designed for the family of hyperfinite graphs. Roughly speaking, hyperfinite graphs are non-expanding graphs. Formally, a graph $G = (V, E)$ is (ϵ, k) -hyperfinite if it is possible to remove at most $\epsilon|V|$ edges of the graph so that the remaining graph has connected components of size at most k . We refer to these edges as the *separating edges* of G . A graph G is ρ -hyperfinite for $\rho : \mathbb{R}_+ \rightarrow \mathbb{R}_+$ if for every $\epsilon \in (0, 1]$, G is $(\epsilon, \rho(\epsilon))$ -hyperfinite. The family of hyperfinite graphs includes many subfamilies of graphs such as graphs with an excluded-minor (e.g. planar graphs), graphs that have subexponential growth and graphs with bounded treewidth. The complexity of our algorithm does not depend on the size of the graph as stated in the next theorem.

Theorem 4 *Algorithm 2, when run with $k = \rho(\epsilon)$, is a local sparse spanning graph algorithm for the family of ρ -hyperfinite graphs with a degree bounded by d . The query complexity and running time of Algorithm 2 are $O(d^{\rho(\epsilon)+1})$, and its success probability is 1.*

We note that we could also obtain a local sparse spanning graph algorithm for hyperfinite graphs by using the partition oracle of [HKNO09] (see the reduction described in Section 6) but the complexity would be higher ($O(d^{d^{\rho(\epsilon)}})$).

We present Algorithm 2 in Subsection 5.1. In Subsection 5.2 we give an improved analysis for the subfamily of graphs that have subexponential growth.

5.1 The Algorithm

Recall that Kruskal's algorithm for finding a minimum-weight spanning tree in a weighted connected graph works as follows. First it sorts the edges of the graph from minimum to maximum weight (breaking ties arbitrarily). Let this order be e_1, \dots, e_m . It then goes over the edges in this order, and adds e_i to the spanning tree if and only if it does not close a cycle

with the previously selected edges. It is well known (and easy to verify), that if the weights of the edges are distinct, then there is a single minimum weight spanning tree in the graph. For an unweighted graph G , consider the order defined over its edges by the order of the ids of the vertices. Namely, we define a ranking r of the edges as follows: $r(u, v) < r(u', v')$ if and only if $\min\{id(u), id(v)\} < \min\{id(u'), id(v')\}$ or $\min\{id(u), id(v)\} = \min\{id(u'), id(v')\}$ and $\max\{id(u), id(v)\} < \max\{id(u'), id(v')\}$. If we run Kruskal's algorithm using the rank r as the weight function (where there is a single ordering of the edges), then we obtain a spanning tree of G .

While the local algorithm we give in this section (Algorithm 2) is based on the aforementioned global algorithm, it does not exactly emulate it, but rather emulates a certain *relaxed* version of it. In particular, it will answer YES for every edge selected by the global algorithm (ensuring connectivity), but may answer YES also on edges not selected by the global algorithm.

Algorithm 2 (Kruskal-based Algorithm)

The algorithm is provided with an integer parameter k , which is fixed for all queries. On query (x, y) :

1. Perform a BFS to depth k from x , thus obtaining the subgraph $C_k(x)$ induced by $\Gamma_k(x)$ in G .
 2. If (x, y) is the edge with largest rank on some cycle in $C_k(x)$, then answer NO, otherwise, answer YES.
-

Proof of Theorem 4: By the description of Algorithm 2 it directly follows that its answers are consistent with a connected subgraph G' . We next show that the algorithm returns YES on at most $(1 + \epsilon)n$ edges. Let $k = \rho(\epsilon)$. For a vertex u , let $\tilde{C}(u) = (\tilde{V}(u), \tilde{E}(u))$ denote the component of u after the removal of the separating edges (as defined at the start of the subsection). We next prove that G' does not contain a cycle on the subgraph induced on $\tilde{V}(u)$. In our proof we use properties of $\tilde{C}(u)$, however, we note that the algorithm does not compute $\tilde{C}(u)$. By definition, $|\tilde{V}(u)| \leq k$, thus the diameter of $\tilde{C}(u)$ is at most $k - 1$. This implies that $C_k(u)$ contains $\tilde{C}(u)$ for every $u \in G$. Let σ be a cycle in $\tilde{C}(u)$ and let $e = (w, v)$ be the edge in σ with the largest rank. Since $\tilde{C}(u) = \tilde{C}(v) = \tilde{C}(w)$ it follows that on query (w, v) the algorithm returns NO. We conclude that for every $u \in V$ the algorithm returns YES only on at most $|\tilde{V}(u)| - 1$ among the edges in $\tilde{E}(u)$. Since the number of edges that do not belong to any component $\tilde{C}(u)$, that is, the number of separating edges in an $(\epsilon, k = \rho(\epsilon))$ -hyperfinite graph is at most $\epsilon|V|$ we have that the total number of edges for which the algorithm returns YES is at most $(1 + \epsilon)|V|$. ■

5.2 Graphs with Subexponential-Growth

In this subsection we analyze Algorithm 2 when executed on graphs with subexponential-growth and for an appropriate k . We first show that graphs with subexponential-growth are $(\epsilon, \rho(\epsilon))$ -hyperfinite. In order to obtain an improved analysis of the complexity of Algorithm 2 for graphs with subexponential-growth, we bound not only the size of each component but also the diameter of each component.

A monotone function $f : \mathbb{N} \rightarrow \mathbb{N}$ has *subexponential growth* if for any $\beta > 0$, there exists $r_f(\beta) > 0$ such that $f(r) \leq \exp(\beta \cdot r)$ for all $r \geq r_f(\beta)$. A graph G has *growth bounded by f* if for every $k \geq 1$, $n_k(G) \leq f(k)$.

Theorem 5 *Given a graph $G = (V, E)$ with degree bounded by d that has growth bounded by $f : \mathbb{N} \rightarrow \mathbb{N}$ where f has subexponential growth, there is a local sparse spanning graph algorithm with query complexity and running time $O(d \cdot n_{r_f(\beta)}(G)) = O(d \cdot \exp(\beta \cdot r_f(\beta)))$ for $\beta = \frac{\epsilon}{2d}$.*

Recall that Algorithm 2 is provided with an integer parameter, k , which determines the depth of the BFS that is performed by the algorithm. In case the graph is $(\epsilon, \rho(\epsilon))$ -hyperfinite we showed that setting $k = \rho(\epsilon)$ is sufficient. For a general graph G , we next define another parameter which is also sufficient for bounding the required depth of the BFS, as we show in Theorem 6. Thereafter, we shall prove that for graphs with subexponential-growth this parameter is small and can be computed efficiently.

Define $k_{\alpha, \beta}^K(G)$ to be the minimum distance k ensuring that all but an α -fraction of the vertices have at most $\exp(\beta k/2)$ vertices in their k -neighborhood (k is allowed to be larger than the diameter of G so that $k_{\alpha, \beta}^K(G)$ is well defined). Formally,

$$k_{\alpha, \beta}^K(G) = \min_k \{ |\{v : |\Gamma_k(v)| \leq \exp(\beta k/2)\}| \geq (1 - \alpha) |V| \} . \quad (3)$$

Theorem 6 *Algorithm 2, when run with $k \geq k_{\alpha, \beta}^K(G)$, where $\alpha + \beta = \epsilon/d$, is a local sparse spanning graph algorithm. The query complexity and running time of the algorithm are $O(dn_k(G)) = O(d^{k+1})$.*

Theorem 5 follows directly from Theorem 6 and Theorem 6 follows directly from the proof of Theorem 4 and the following lemma.

Lemma 1 *Every graph $G = (V, E)$ is $(\epsilon, (1 + \beta)^k)$ -hyperfinite for $k = k_{\alpha, \beta}^K(G)$ and $\alpha + \beta = \epsilon/d$. Moreover, it is possible to remove at most $\epsilon|V|$ edges of the graph so that the remaining graph has connected components with diameter at most $2k$.*

Proof: Let $S \subseteq V$ denote the set of vertices, v , for which $|\Gamma_k(v)| > \exp(\beta k/2)$. We start by removing all the edges adjacent to vertices in S . Overall, we remove at most $d\alpha|V|$ edges. For each vertex $v \in V - S$ it holds that $|\Gamma_k(v)| \leq \exp(\beta k/2)$. From the fact that $\exp(x) < 1 + 2x$ for every $x < 1$ we obtain that $|\Gamma_k(v)| < (1 + \beta)^k$. Therefore, there exists

$k' < k$ such that $|\Gamma_{k'+1}(v)| < |\Gamma_{k'}(v)|(1 + \beta)$. Thus, $C_{k'}(v)$ can be disconnected from G by removing at most $d\beta|\Gamma_{k'}(v)|$ edges. Since it holds that $|\Gamma_k(v)| < (1 + \beta)^k$ for every subgraph of G and every $v \in V - S$, we can continue to iteratively disconnect connected components of diameter at most $2k$ from the resulting graph. Hence, we obtain that by removing at most $d(\alpha + \beta)|V|$ edges, the remaining graph has connected components with diameter at most $2k$, as desired. ■

5.3 The Parameter k

Recall that Algorithm 2 is given a parameter k that determines the depth of the BFS that the algorithm performs. By Theorem 6 it is sufficient to require that $k \geq k_{\epsilon/(2d), \epsilon/(2d)}^K(G)$ in order to ensure that the resulting graph is sparse. For the case that k is not given in advance, we can compute k such that with probability greater than $1 - \delta$ it holds that

$$k_{\epsilon/(2d), \epsilon/(2d)}^K(G) \leq k \leq k_{\epsilon/(4d), \epsilon/(2d)}^K(G), \quad (4)$$

as follows. Sample $\Theta(1/\epsilon^2 \log(1/\delta))$ vertices. Start with $k = 1$ and iteratively increase k until for at least $(1 - \frac{3\epsilon}{8d})$ -fraction of the vertices, v , in the sample it holds that $|\Gamma_k(v)| \leq \exp(\epsilon k/(4d))$. By Chernoff's inequality we obtain that with probability greater than $1 - \delta$ Equation (4) holds.

6 Partition Oracle Based Algorithm

In this section we describe a simple reduction from local algorithm for sparse spanning graph to partition oracle. We begin with a few definitions concerning partition oracles.

Definition 3 For $\epsilon \in (0, 1]$, $k \geq 1$ and a graph $G = (V, E)$, we say that a partition $\mathcal{P} = (V_1, \dots, V_t)$ of V is an (ϵ, k) -partition (with respect to G), if the following conditions hold:

1. For every $1 \leq i \leq t$ it holds that $|V_i| \leq k$;
2. For every $1 \leq i \leq t$ the subgraph induced by V_i in G is connected;
3. The total number of edges whose endpoints are in different parts of the partition is at most $\epsilon|V|$ (that is, $|\{(v_i, v_j) \in E : v_i \in V_j, v_j \in V_j, i \neq j\}| \leq \epsilon|V|$).

Let $G = (V, E)$ be a graph and let \mathcal{P} be a partition of V . We denote by $g_{\mathcal{P}}$ the function from $v \in V$ to 2^V (the set of all subsets of V), that on input $v \in V$, returns the subset $V_{\ell} \in \mathcal{P}$ such that $v \in V_{\ell}$.

Definition 4 ([HKNO09]) An oracle \mathcal{O} is a partition oracle if, given query access to the incidence-lists representation of a graph $G = (V, E)$, the oracle \mathcal{O} provides query access to a partition $\mathcal{P} = (V_1, \dots, V_t)$ of V , where \mathcal{P} is determined by G and the internal randomness

of the oracle. Namely, on input $v \in V$, the oracle returns $g_{\mathcal{P}}(v)$ and for any sequence of queries, \mathcal{O} answers consistently with the same \mathcal{P} . An oracle \mathcal{O} is an (ϵ, k) -partition oracle with respect to a class of graphs \mathcal{C} if the partition \mathcal{P} it answers according to has the following properties.

1. For every $V_\ell \in \mathcal{P}$, $|V_\ell| \leq k$ and the subgraph induced by V_ℓ in G is connected.
2. If G belongs to \mathcal{C} , then $|\{(u, v) \in E : g_{\mathcal{P}}(v) \neq g_{\mathcal{P}}(u)\}| \leq \epsilon|V|$ with high constant probability, where the probability is taken over the internal coin flips of \mathcal{O} .

By the above definition, if $G \in \mathcal{C}$, then with high constant probability the partition \mathcal{P} is an (ϵ, k) -partition, while if $G \notin \mathcal{C}$ then it is only required that each part of the partition is connected and has size at most k .

Theorem 7 *If there exists an (ϵ, k) -partition oracle, \mathcal{O} , for the family of graphs \mathcal{C} having query complexity $q(\epsilon, k, d, n)$ and running time $t(\epsilon, k, d, n)$, then there exists a local sparse spanning graph algorithm, \mathcal{A} , for the family of graphs \mathcal{C} , whose success probability is the same as that of \mathcal{O} . The running time of \mathcal{A} is bounded from above by $t(\epsilon, k, d, n) + O(kd)$ and the query complexity of \mathcal{A} is $q(\epsilon, k, d, n) + O(kd)$.*

Proof: On query (u, v) the algorithm \mathcal{A} proceeds as follows:

1. Query \mathcal{O} on u and v and get $g_{\mathcal{P}}(u)$ and $g_{\mathcal{P}}(v)$, respectively.
2. If $g_{\mathcal{P}}(u) \neq g_{\mathcal{P}}(v)$, return YES.
3. Otherwise, let w denote the vertex in $g_{\mathcal{P}}(u)$ such that $id(w)$ is minimal.
4. Perform a BFS on the subgraph induced on $g_{\mathcal{P}}(u)$, starting from w .
5. If (u, v) belongs to the edges of the above BFS then return YES, otherwise, return NO.

The fact that \mathcal{A} returns YES on at most $(1 + \epsilon)|V|$ edges follows from the fact that \mathcal{P} is a partition can that $|\{(u, v) \in E : g_{\mathcal{P}}(v) \neq g_{\mathcal{P}}(u)\}| \leq \epsilon|V|$. The connectivity follows from the fact that the subgraph induced on V_i is connected for every $V_i \in \mathcal{P}$. The additional term of $O(kd)$ in the time and query complexity is due to the BFS performed on $g_{\mathcal{P}}(u)$. ■

The following corollaries follow from [EHNO11] and [LR13a, LR13b], respectively.

Corollary 8 *There exists a local sparse spanning graph algorithm for the family of graphs with bounded treewidth. This algorithm has high constant success probability and its query complexity and running time are $\text{poly}(1/\epsilon, d)$.*

Corollary 9 *There exists a local sparse spanning graph algorithm for the family of graphs with a fixed excluded minor. This algorithm has high constant success probability and its query complexity and running time are $(d/\epsilon)^{O(\log(1/\epsilon))}$.*

7 Graphs with an Excluded Minor - the Weighted Case

In this section we focus on another subfamily of hyperfinite graphs, graphs with an excluded fixed minor. We consider a more general setting of the problem in which the edges of the graph are weighted and the goal is to locally find a spanning graph with small weight (see Definition 5). The time and query complexity of the local algorithm we provide is quasi-polynomial in $1/\epsilon$ and d .

For the unweighted case, we note that by a reduction to partition oracle for graphs with an excluded minor (see Section 6) we can obtain a local algorithm for sparse spanning graph that also have a quasi-polynomial complexity ⁶.

A graph H is called a *minor* of a graph G if H is isomorphic to a graph that can be obtained by zero or more edge contractions on a subgraph of G . A graph G is *H -minor free* or *excludes H as a minor* if H is not a minor of G .

Definition 5 *A local algorithm for $(1 + \epsilon)$ -approximating the minimum weight spanning graph of a graph $G = (V, E, w)$ with positive weights and $\min_{e \in E} w(e) \geq 1$, is a local algorithm for $(1 + \epsilon)$ -sparse spanning graph of $G = (V, E, w)$ for which the following holds: $\sum_{e \in E'} w(e) \leq (1 + \epsilon)\alpha$, where α is the weight of a minimum weight spanning tree of G .*

For a graph $G = G = (V, E, w)$ define $W_G = \max_{e \in E} w(e)$ (for the sake of presentation, when it is clear from the context, we sometimes omit the subscript G). We shall prove the following theorem:

Theorem 10 *There exists a local algorithm for $(1 + \epsilon)$ -approximating the minimum weight spanning graph for the family of graphs with a fixed excluded minor, with high constant success probability and time and query complexity $(Wd/\epsilon)^{O(\log(W/\epsilon))}$.*

We note that the complexity of the algorithm must depend on the maximum edge weight, W . To see why this is true, consider the following families of graphs. In the first family, each graph consists of $s = \Theta(n/W)$ cycles of size $\Theta(W)$: $\mathcal{C}_1, \dots, \mathcal{C}_s$. Each cycle \mathcal{C}_i is connected to \mathcal{C}_{i+1} with a single edge and in each \mathcal{C}_i there is a single edge of weight W and all the other edges have weight 1. We similarly define the second family with the exception that we remove a random edge from each cycle. It is easy to see that in order to distinguish between these two families we need to make $\Omega(W)$ queries.

We begin with describing a global algorithm that iteratively selects edges where all the edges selected by the algorithm belong to a minimum weight spanning tree (Subsection 7.1). Thereafter, we describe a local algorithm that by local simulation selects the same edges that the global algorithm selects along with an additional small set of edges (Subsection 7.2).

⁶At a high level the technique in our algorithm appears similar to the one in [LR13a]. However, due to the difference in the operation of contraction (Definition 6) the behavior of the algorithm is very different.

7.1 The Global Algorithm

We begin by describing a global partitioning algorithm. In the description of our global partitioning algorithm we shall use the following definitions and subroutines.

We first define the operation of *contraction*. Roughly speaking, given a partition of the graph into connected components, the operation of contraction replaces each connected component with a single vertex. A pair of vertices in the new graph are connected with an edge if the edge cut of the corresponding connected components is not empty. The weight of the edge is set to be the weight of the edge with the minimum weight in the edge cut, as explained formally in what follows.

For a graph $G = (V, E)$ and two sets of vertices $V_1, V_2 \subseteq V$, we let $E(V_1, V_2)$ denote the set of edges in G with one endpoint in V_1 and one endpoint in V_2 . That is $E(V_1, V_2) \stackrel{\text{def}}{=} \{(v_1, v_2) \in E : v_1 \in V_1, v_2 \in V_2\}$.

Definition 6 *Let $G = (V, E, w)$ be an edge-weighted graph and let $\mathcal{P} = (V_1, \dots, V_t)$ be a partition of the vertices of G such that for every $1 \leq i \leq t$, the subgraph induced by V_i is connected. Define the contraction G/\mathcal{P} of G with respect to the partition \mathcal{P} to be the edge-weighted graph $G' = (V', E', w')$ where:*

1. $V' = \{V_1, \dots, V_t\}$ (that is, there is a vertex in V' for each subset of the partition \mathcal{P});
2. $(V_i, V_j) \in E'$ if and only if $i \neq j$ and $E(V_i, V_j) \neq \emptyset$;
3. $w'((V_i, V_j)) = \min_{(u,v) \in E(V_i, V_j)} w((u, v))$.

We note that in the definition of contraction that appears in [LR13a] the weight of an edge in the new graph is set to be the sum of weights of the edges in the cut as opposed to the minimum weight as it appears in Item 3 of Definition 6.

As a special case of Definition 6 we get the standard notion of a single-edge contraction.

Definition 7 *Let $G = (V, E, w)$ be an edge-weighted graph on n vertices v_1, \dots, v_n , and let (v_i, v_j) be an edge of G . The graph obtained from G by contracting the edge (v_i, v_j) is G/\mathcal{P} where \mathcal{P} is the partition of V into $\{v_i, v_j\}$ and singletons $\{v_k\}$ for every $k \neq i, j$.*

We assume without loss of generality that a graph $G = (V, E, w)$ has distinct weights and thus a unique minimum weight spanning tree denoted by $MST(G)$. This can be achieved by determining that for every $e' \in E$, $e \in E$ for which $w(e) = w(e')$ we have $w(e) > w(e')$, where e is the edge with higher rank amongst e' and e .

Mader [Mad67] proved that a sufficiently large average degree guarantees a K_t -minor as stated in the next fact.

Fact 1 *Let H be a fixed graph. There is a constant $c_1(|H|)$, such that in every H -minor free graph, $G = (V, E)$, it holds that $|E| \leq c_1(|H|) \cdot |V|$.*

In our algorithm we shall use the following subroutine.

Theorem 11 ([LR13a], Corollary 2) *Let H be a fixed graph. There is a constant $c_2(H) > 1$ such that for every $\gamma \in (0, 1]$, every H -minor free graph $G = (V, E)$ with degree bounded by d is $(\gamma, c_2(H)d^2/\gamma^2)$ -hyperfinite. Furthermore, there exists a set S of vertices such that:*

1. $|S| = (\gamma/d)|V|$.
2. The removal of S leaves G with connected components of size at most $c_2(H)d^2/\gamma^2$ each.
3. S can be found in time $O(|V|^{3/2})$.

Algorithm 3 A global partition algorithm for an H -minor free graph $G = (V, E)$

1. Set $G^0 := G$
 2. For $i = 1$ to $\ell = \Theta(\log W/\epsilon)$:
 - (a) Toss a fair coin for every vertex in G^{i-1} .
 - (b) For each vertex u let (u, v) be an edge with minimum weight that is incident to u (where ties are broken according to the rank of the edges). If u 's coin toss is 'Heads' and v 's coin toss is 'Tails', then contract (u, v) .
 - (c) Let $\tilde{G}^i = (\tilde{V}^i, \tilde{E}^i, \tilde{w}^i)$ denote the graph resulting from the contraction of the edges as determined in the previous step. Hence, each vertex $\tilde{v}_j^i \in \tilde{V}^i$ corresponds to a subset of vertices in G , which we denote by \tilde{C}_j^i .
 - (d) Let $\gamma = \epsilon/(6W\ell)$. For each \tilde{C}_j^i such that $|\tilde{C}_j^i| > c_2(H)d^2/\gamma^2$, partition the vertices in \tilde{C}_j^i into connected subsets of size at most $k = c_2(H)d^2/\gamma^2$ each by running the algorithm referred to in Theorem 11 on the subgraph induced by \tilde{C}_j^i in G .
 - (e) Set $G^i := G/\mathcal{P}^i$, where \mathcal{P}^i is the partition resulting from the previous step.
 3. For each subset C_j^ℓ in \mathcal{P}^ℓ such that $|C_j^\ell| > c_2(H)W^2d^2/\epsilon^2$, partition the vertices in C_j^ℓ into connected subsets each of size at most $3c_2(H)W^2d^2/\epsilon^2$ by running the algorithm referred to in Theorem 11, and output the resulting partition.
-

We begin by describing a global partitioning algorithm with the following properties.

Theorem 12 *Let H be a fixed graph. If the input graph $G = (V, E, w)$ is H -minor free and has degree bounded by d , then for any given $\epsilon \in (0, 1]$, with high constant probability, Algorithm 3 outputs a partition $\mathcal{P} = \{V_1, \dots, V_m\}$ of G for which the following holds:*

1. The subgraph induced on V_i is connected for every $i \in [m]$.

$$2. |V_i| = O(W^2 d^2 / \epsilon^2).$$

$$3. \sum_{e \in F} w(e) \leq \epsilon |V| \text{ where } F = \{(u, v) \in E : \exists i \neq j \text{ s.t. } u \in V_i \wedge v \in V_j\}.$$

In addition every edge that Algorithm 3 contracts belongs to $MST(G)$.

Proof: The fact that every edge that Algorithm 3 contracts belongs to $MST(G)$ follows from the fact that the minimum-weight edge adjacent to every vertex belongs to the minimum weight spanning tree of G . Items (1)-(2) of the theorem follow from the construction. Hence, in what follows we prove item (3). We say that an iteration i is *successful* if $|\tilde{V}^i| \leq 15|V^{i-1}|/16$. For every $v \in V^{i-1}$, Define I_v to be the indicator variable for the event that an edge selected by v is contracted. Recall that this event occurs if the coin flip of v is ‘Heads’ and that of u is ‘Tails’, implying that $\Pr[I_v] = 1/4$. Note that for any outcome of coin flips it holds that $|\tilde{V}^i| = |V^{i-1}| - \sum_{v \in V^{i-1}} I_v$. Thus we obtain that $E(|\tilde{V}^i|) = 3|V^{i-1}|/4$. By Markov’s inequality, the probability of success is at least $\eta = 4/5$, for every iteration ⁷. Let Y_i be the random variable that takes the value $-1/\eta + 1$ if the i -th iteration is successful and takes the value 1 otherwise. Let $Z_i = \sum_{j=1}^i Y_j$, then $E(Z_{k+1}|Z_1, \dots, Z_k) \leq Z_k + 1 \cdot (1 - \eta) + (1 - 1/\eta) \cdot \eta = Z_k$ and $|Z_k - Z_{k-1}| \leq \frac{1}{\eta}$ for every k . By Azuma’s inequality [Azu67] we obtain that $\Pr(Z_\ell \geq t) \leq e^{-t^2/(2\ell(1/\eta)^2)}$ for every t . Setting $t = 3\sqrt{\ell}/\eta$, we obtain that with probability greater than 9/10, $Z_\ell < 3\sqrt{\ell}/\eta$. Let s denote the number of successful iterations, then $Z_\ell = (\ell - s) + s(1 - 1/\eta) = \ell - s/\eta$. Thus, we obtain that $s > \eta\ell - 3\sqrt{\ell}$. We conclude that with probability at least 9/10, the number of successful iterations is at least $\eta\ell/2 = 2\ell/5$ for sufficiently large ℓ .

Our second claim is that for every i , after Step 2d, it holds that $|V^i| \leq |\tilde{V}^i| + \frac{\epsilon n}{3W\ell}$ (where $n = |V|$). This follows from Theorem 11: For each component \tilde{C}_j^i that we break, we increase the total number of vertices by an additive term of at most $2\gamma|\tilde{C}_j^i| = \frac{\epsilon|\tilde{C}_j^i|}{3W\ell}$. Thus, after Step 2d of the ℓ^{th} iteration, with probability at least 9/10, the number of vertices in G^ℓ is at most

$$c_1(H) \cdot n \cdot \left(1 - \frac{1}{16}\right)^{\frac{2\ell}{5}} + \ell \cdot \frac{\epsilon n}{3W\ell} \leq \frac{2\epsilon n}{3W}. \quad (5)$$

Since we add at most $\epsilon n/(3W)$ vertices in Step 3 (when breaking the subsets corresponding to vertices in G^ℓ into subsets of size at most $3c_2(H)W^2 d^2/\epsilon^2$), we obtain the desired result. \blacksquare

7.2 The Local Algorithm

In what follows we prove Theorem 10. The proof consists of two parts. In the first part we describe how, given query access to the incidence-lists representation of a graph $G = (V, E, w)$

⁷We note that we only use the independence of the coin tosses between adjacent vertices. Therefore, in the i -th iteration it is enough to require $2d_i$ -wise independence, where d_i denotes the maximum degree of G^{i-1} .

and a vertex $v \in V$, it is possible to simulate Algorithm 3 locally and determine the part that v belongs to in the partition \mathcal{P} that the algorithm outputs. In the second part of the proof we describe how to obtain a local algorithm for approximating the minimum weight spanning graph, given query access to the above partition.

Proof of Theorem 10: Recall that the partition Algorithm 3 outputs, \mathcal{P} , is determined randomly based on the ‘Heads’/‘Tails’ coin-flips of the vertices in each iteration. Also recall that $g_{\mathcal{P}}(v)$ denotes the subset of vertices that v belongs to in \mathcal{P} . Since we want the algorithm to be efficient, the algorithm will flip coins “on the fly” as required for determining $g_{\mathcal{P}}(v)$. Since the algorithm has to be consistent with the same \mathcal{P} for any sequence of queries it gets, it will keep in its memory all the outcomes of the coin flips it has made. For the sake of simplicity, whenever an outcome of a coin is required, we shall say that a coin is flipped, without explicitly stating that first the algorithm checks whether the outcome of this coin flip has already been determined.

Recall that the algorithm constructs a sequence of graphs $G^0 = G, \tilde{G}^1, G^1, \dots, \tilde{G}^\ell, G^\ell$, and that for each $0 \leq i \leq \ell$, the vertices in G^i correspond to connected subgraphs of G (which we refer to as components). For a vertex $v \in V$ let $C^i(v)$ denote the vertex/component that v it belongs to in G^i , and define $\tilde{C}^i(v)$ analogously with respect to \tilde{G}^i . Indeed, we shall refer to vertices in G^i (\tilde{G}^i) and to the components that correspond to them, interchangeably. When the algorithm flips a coin for a vertex C in G^i , we may think of the coin flip as being associated with the vertex having the largest id in the corresponding component in G .

Let $Q^i(v)$ denote the number of queries to G that are performed in order to determine $C^i(v)$, and let Q^i denote an upper bound on $Q^i(v)$ that holds for any vertex v . We first observe that $Q^1 \leq d^2$. In order to determine $C^1(v)$, the algorithm first flips a coin for v . If the outcome is ‘Tails’ then the algorithm queries the neighbors of v . For each neighbor u of v it determines whether (u, v) is the lightest edge incident to u (by querying all of u ’s neighbors). If so, it flips a coin for u , and if the outcome is ‘Heads’, then the edge is contracted (implying that $u \in \tilde{C}^1(v)$). If v is a ‘Heads’ vertex, then it finds its lightest incident edge, (v, u) by querying all of v ’s neighbors. If u is a ‘Tails’ vertex (so that (v, u) is contracted), then the algorithm queries all of u neighbors, and for each neighbor it queries all of its neighbors. By doing so (and flipping all necessary coins) it can determine which additional edges (u, y) incident to u are contracted (implying for each that $y \in \tilde{C}^1(v) = \tilde{C}^1(u)$). In both cases (of the outcome of v ’s coin flip), the number of queries performed to G is at most d^2 . Recall that a component as constructed above is ‘broken’ if it contains more than $k = \tilde{O}(W^2 d^2 / \epsilon^2)$ vertices. Since $|\tilde{C}^1(v)| \leq d + 1$ for every v , we have that $C^1(v) = \tilde{C}^1(v)$.

For general $i > 1$, to determine the connected component that a vertex v belongs to after iteration i , we do the following. First we determine the component it belongs to after iteration $i - 1$, namely $C^{i-1}(v)$, at a cost of at most Q^{i-1} queries. Note that by the definition of the algorithm, $|C^{i-1}(v)| \leq k$. We now have two cases:

Case 1: $C^{i-1}(v)$ is a ‘Tails’ vertex for iteration i . In this case we query all edges incident to vertices in $C^{i-1}(v)$, which amounts to at most $d \cdot k$ edges. For each endpoint u of such an edge we find $C^{i-1}(u)$. For each $C^{i-1}(u)$ that is ‘Heads’ we determine whether its lightest

incident edge connects to $C^{i-1}(v)$ and if so the edge is contracted (so that $C^{i-1}(u) \subset \tilde{C}^i(v)$). To do so, we need, again, to query all the edges incident to vertices in $C^{i-1}(u)$. The total number of vertices x for which we need to find $C^{i-1}(x)$ is upper bounded by $dk + 1$. The total number of additional queries we make is upper bounded by d^2k^2 .

Case 2: $C^{i-1}(v)$ is a ‘Heads’ vertex in iteration i . In this case we find its lightest incident edge in G^{i-1} and the other endpoint in G^{i-1} . Let C' denote the other endpoint in G^{i-1} . If C' is a ‘Tails’ vertex then we apply the same procedure to C' as described in Case 1 for $C^{i-1}(v)$ (that is, in the case that $C^{i-1}(v)$ is a ‘Tails’ vertex in G^{i-1}). The bound on the number of queries performed is also as in Case 1. In either of the two cases we might need to ‘break’ $\tilde{C}^i(v)$ (in case $|\tilde{C}^i(v)| > k$) so as to obtain $C^i(v)$. However, this does not require performing any additional queries to G since all edges between vertices in $\tilde{C}^i(v)$ are known, and this step only contributes to the running time of the algorithm. We thus get the following recurrence relation for Q^i : $Q^i = d^2 \cdot k^2 + (d \cdot k + 1) \cdot Q^{i-1}$. Since $k = \text{poly}(Wd/\epsilon)$ we get that

$$Q^\ell \leq \ell \cdot d^2 \cdot k^2 + (d \cdot \text{poly}(Wd/\epsilon))^\ell = (Wd/\epsilon)^{O(\log(W/\epsilon))}, \quad (6)$$

as claimed.

We next turn to bounding the running time. Let $T^i(v)$ denote the running time for determining $C^i(v)$. By the same reasoning as above we have that $T^i \leq O(d \cdot k) \cdot T^{i-1} + B$ where B is an upper bound on the running time of breaking a connected component at each iteration. From Theorem 11 we obtain that $B \leq (d \cdot k^2)^{3/2}$. Thus, the running time of the algorithm is $(Wd/\epsilon)^{O(\log(W/\epsilon))}$ for a single query. As explained above, for the sake of consistency, the algorithm stores its previous coin-flips. By using a balanced search tree to store the coin flips we obtain that the total running time of the algorithm for a sequence of q queries is $q \log q \cdot (Wd/\epsilon)^{O(\log(W/\epsilon))}$, as claimed.

Finally, we turn to describing how, by locally simulating Algorithm 3 we obtain a local algorithm for approximating the minimum weight spanning graph. On query (u, v) simulate Algorithm 3 on query u and query v . If $g_{\mathcal{P}}(u) \neq g_{\mathcal{P}}(v)$, return YES if and only if (u, v) is that lightest edge which has one endpoint in $g_{\mathcal{P}}(u)$ and one end point in $g_{\mathcal{P}}(v)$. Otherwise, return YES if and only if (u, v) was contracted, at some point, during the local simulation of Algorithm 3.

Let E' denote the set of edges for which the local algorithm returns a positive response. From that fact that every edge that Algorithm 3 contracts belongs to $MST(G)$ and the fact that the number of edges (u, v) for which $g_{\mathcal{P}}(u) \neq g_{\mathcal{P}}(v)$ is bounded by $\epsilon n/W$ we conclude that $\sum_{e \in E'} w(e) \leq \alpha + \epsilon|V|$ as desired. We conclude the proof by showing that the subgraph $G' = (V, E')$ is connected. Observe that E' contains all the edges contracted by Algorithm 3. Thus, the subgraph induced in G on $g_{\mathcal{P}}(u)$ is connected for every u . In addition for every u and v for which $g_{\mathcal{P}}(u) \neq g_{\mathcal{P}}(v)$, if the subset of edges of G , $S(u, v)$, with one endpoint in $g_{\mathcal{P}}(u)$ and another endpoint in $g_{\mathcal{P}}(v)$, is not empty, then E' contains at least one edge of $S(u, v)$, namely $E' \cap S(u, v) \neq \emptyset$. We conclude that G' is connected. Since both the time complexity and query complexity of the local algorithm is dominated by the local simulation complexity of Algorithm 3 we derive the same bounds on the corresponding complexities. ■

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