

On Sparse Spanners of Weighted Graphs*

Hugo Althöfer,¹ Gautam Das,² David Dobkin,³ Deborah Joseph,⁴ and José Soares⁵

¹ Fakultät für Mathematik, Universität Bielefeld, Postfach 8640,
4800 Bielefeld, Federal Republic of Germany

² Mathematical Sciences Department, Memphis State University,
Memphis, TN 38152, USA
dgd@ncsl.mcsimems.tn.edu

³ Department of Computer Science, Princeton University,
Princeton, NJ 08544, USA
dgd@princeton.edu

⁴ Department of Computer Sciences, University of Wisconsin,
Madison, WI 53706, USA
joseph@cs.wisc.edu

⁵ Universidade de São Paulo, IME-USP/MAC,
R. do São Paulo-SP, Brazil, and
Department of Computer Science, University of Chicago,
Chicago, IL 60637, USA
soares@uchicago.edu

Abstract. Given a graph G , a subgraph G' is a t -spanner of G if, for every $u, v \in V$, distance from u to v in G' is at most t times longer than the distance in G . In this paper we give a simple algorithm for constructing sparse spanners for arbitrary weighted graphs. We then apply this algorithm to obtain specific results for planar graphs and Euclidean graphs. We discuss the optimality of our results and present several nearly matching lower bounds.

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with stretch factor $O(t)$, and weight within an $O(1 + 1/t)$ multiple of the weight of the minimum spanning tree. Weight sparseness has also been considered in [AJ], [D], and [S] for the special case of 1-spanners under a general model where spanners may have auxiliary vertices.

In this paper we approach the problem from a very broad perspective. Our graphs have no special embeddings, and we allow arbitrary positive edge weights. For any t , we show that every such graph has a spanner with an $O(t)$ stretch factor, and $O(1 + 1/t)$ size. We also provide weight bounds for our spanners. The contribution of this paper are: a very simple polynomial-time algorithm for constructing sparse spanners (sparse both in size and weight) of arbitrary weighted graphs, using these ideas for constructing spanners of planar graphs, some lower bound results, and some results on spanners in Euclidean spaces of arbitrary dimensions and norms. Since any spanner with appropriate sparseness and stretch factor can be used for constructing synchronizers [PUI], our algorithm provides a simple alternative to the clustering algorithm previously used for constructing synchronizers. Similarly, we hope that our results will simplify the construction of succinct routing tables for arbitrary weighted graphs.

2. Construction of Sparse Spanners

Let G be an n -vertex, connected, weighted graph. Let us examine the minimum spanning tree, denoted as $MST(G)$, as a possible candidate for a spanner. It is obviously the sparsest spanner under both criteria of sparseness, because it is the subgraph with the least total edge weight that still connects all vertices. However, its stretch factor can be as bad as $\Omega(n)$. To see this, consider an n -vertex spanning tree does not contain an edge, say $[u, v]$, and the distance from u to v in the tree is $n - 1$. Thus the tree is an $(n - 1)$ -spanner.

Instead, we would like to look for spanners whose stretch factors are constants, independent of n . We use the minimum spanning tree as a yardstick for measuring sparseness, and try to generate spanners whose sparseness comes close to it. Our results are encouraging, because we show that any graph has spanners with constant stretch factors, whose sparseness can be made arbitrarily close to that of $MST(G)$. They are summarized by Theorems 1 and 2.

Theorem 1. Given an n -vertex graph G and a $t > 0$, there is a polynomially constructible $(2t + 1)$ -spanner G' such that,

- (1) $Size(G') < n^{1/t}$
- (2) $Weight(G') < Weight(MST(G))(1 + n/2t)$

Note that the stretch factor is independent of the number of vertices in the graph, and of the edge weights. Thus, even for a dense graph with $\Omega(n^2)$ edges, sparse spanners exist which have good stretch factors. Furthermore, the sparseness can be made arbitrarily close to that of $MST(G)$, at the expense of an increase in the stretch factor. Previous research [Aw], [PS] has produced spanners with the same size and stretch factors, but the graphs considered there were unweighted and sparseness was measured by size only. Also, our algorithm is simpler than the clustering algorithm employed. These bounds should be compared with the following lower bounds, to be proved in Section 3:

Theorem 4. For every pair of integers $t \geq 0$ and $n \geq 3$, there exists an n -vertex graph G for which every $(2t + 1)$ -spanner G' is such that,

$$(1) \text{Size}(G') > \frac{8}{3} n^{1+4/(2t+3)},$$

$$(2) \text{Weight}(G') > \frac{8}{3} \text{Weight}(MST(G))^{n^{4/(2t+3)}}$$

Actually, a slightly weaker lower bound holds even for Steiner spanners, where auxiliary vertices are allowed to be added to the graph (see Section 4). The next theorem concerns planar graphs.

Theorem 2. Given an n -vertex planar graph H and a $t > 0$, there is a polynomially constructible $(2t + 1)$ -spanner G' such that,

$$(1) \text{Size}(G') < (n - 1)(1 + 1/t),$$

$$(2) \text{Weight}(G') < \text{Weight}(MST(G))^{(1+1/t)^2}$$

Here we again observe that the stretch factor is independent of the number of vertices in the planar graph, and of the edge weights. Thus, arbitrarily sparse spanners exist which have good stretch factors. It is well known that the maximum size of a planar graph can be $3n - 6$, so in effect we demonstrate that even the multiplicative constant in the size can be reduced. Theorem 2 is stronger in favor than Theorem 1, because our lower bound results will show that both size and weight bounds are tight. We later give an interesting application of this in connection with Euclidean graphs.

Before we prove our results, we introduce an algorithm for constructing spanners. It constructs a sparse subset of edges so that a required stretch factor is achieved. The algorithm, called $SPANNER(G, r)$, takes as input a weighted graph G , and a positive parameter r . The weights need not be unique. It produces as output a subgraph G' . Note that this algorithm is essentially a generalization of Kruskal's algorithm for computing minimum spanning trees [T]. The algorithm has also been independently discovered by Bern [Be] in a different context, in connection with the problem of dilation of metric spaces [RS]. While it is very simple and easy to implement in polynomial time, the subgraph it generates has many interesting properties.

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Algorithm SPANNER
 Sort E by nondecreasing
 $G := (V, \{\});$
 for every edge $e = [u, v]$
 begin
 Compute $P(u, v)$, the
 if $(r \cdot \text{Weight}(e) < W$
 add e to G' ;
 end;
 Output G' ;
 end;

Lemma 1. G' is an r -spanner of G .
Proof. Consider any edge $[u, v]$ in the following is true. At the instant path $P(u, v)$ from u to v in the current final output each deleted edge in the shortest path between u and v is clearly in G' , and has length

Lemma 2. Let C be any cycle in G .
Proof. Assume that a cycle C remains in the cycle with the largest edges in C to be examined by the algorithm, the remaining portion of G . This portion has at most r edges in C . Thus, at that instant $\text{Weight}([u, v]) \leq r \cdot \text{Weight}([u, v])$. In that case contradiction.

Lemma 3. Let C be any cycle $\text{Weight}(C) - \{e\} > r \cdot \text{Weight}(e)$.
Proof. Assume that a cycle C remains in C . Clearly, it will be examined by the algorithm, say as large as the weight of e . As in many) edges in the cycle with the

Algorithm SPANNER($G = (V, E), r$)

Sort E by nondecreasing weight;

$G := (V, \{\});$

for every edge $e = [u, v]$ in E do

begin

Compute $P(u, v)$, the shortest path from u to v in G ;

if $(r \cdot \text{Weight}(e) < \text{Weight}(P(u, v)))$ then

add e to G ;

end;

Output G ;

end;

Following lemmas describe the properties of the output graph G .

Lemma 1. G is an r -spanner of G .

Consider any edge $[u, v]$ in $G - G$. Since it was rejected by the algorithm, following is true. At the instant it is examined by the algorithm, there is a path $P(u, v)$ from u to v in the current graph of length $\leq r \cdot \text{Weight}([u, v])$. Thus, the final output each deleted edge $[u, v]$ is associated with a short path $P(u, v)$ now consider any shortest path between vertices a and b in G . Let the length of this path be l . For every edge $[u, v]$ along this path not in G , we replace the edge by the corresponding short path $P(u, v)$. The resulting path between a and b is clearly in G , and has length $\leq l \cdot r$, which proves the lemma.

□

Lemma 2. Let C be any cycle in G . Then $\text{Size}(C) > r + 1$.

Proof. Assume that a cycle C remained with size $\leq r + 1$. Let $[u, v]$ be the last edge in C to be examined by the algorithm. Clearly, it is one of the (possibly many) edges in the cycle with the largest weight. When $[u, v]$ is being considered by the algorithm, the remaining portion of the cycle from u to v has already been included in G . This portion has at most r edges, and each edge has weight not greater than $\text{Weight}([u, v])$. Thus, at that instant there is a path from u to v in G , whose length is $\leq r \cdot \text{Weight}([u, v])$. In that case $[u, v]$ should not have been added, which is a contradiction.

Lemma 3. Let C be any cycle in G , and let e be any edge in C . Then $\text{Weight}(C) - \text{Weight}(e) > r \cdot \text{Weight}(e)$.

Proof. Assume that a cycle C remained which violates the above weight condition for some edge e in C . Clearly, it violates this condition for the last edge in C to be examined by the algorithm, say $[u, v]$, because the weight of $[u, v]$ is at least as large as the weight of e . As in the above lemma, $[u, v]$ is one of the (possibly many) edges in the cycle with the largest weight. Thus, the remaining portion of

the cycle has length $> r \cdot \text{Height}([u, v])$. In this case $[u, v]$ should not have been added, which is a contradiction. \square

Lemma 4. $MST(G)$ is contained in G' .

Proof. Before we prove this lemma, let us repeat that our algorithm is essentially a generalized minimum spanning tree algorithm. In fact, for an infinite r its behavior is exactly like Kruskal's minimum spanning tree algorithm [T].

Our proof is by induction on the order in which edges are examined. Let the sequence $\{G_0, G_1, \dots, G_{\text{Size}(G)}\}$ represent the growth of G , where G_i represents the partially constructed subgraph after the i th edge has been examined. At any stage the subgraph will be a collection of connected components, which will finally become one connected component.

Now let us consider Kruskal's algorithm, which also examines edges by nondecreasing weight. Let the sequence $\{M_0, M_1, \dots, M_{\text{Size}(G)} = MST(G)\}$ represent the growth of the minimum spanning tree, where M_i represents the partially constructed tree after the i th edge has been examined. This edge is added only if it does not form a cycle in the partially constructed tree. Thus at any stage, the partially constructed tree will be a forest, which will finally become one connected tree.

We now prove by induction the following, which is a stronger hypothesis than the lemma. For all i , the number of connected components of M_i is the same as that of G_i and each component of M_i is contained in a corresponding component of G_i .

To prove this, let us imagine that both algorithms are being run simultaneously. The basis of the induction is easily proved, because at the start both the tree and the spanner have n components each, consisting of isolated vertices. For the induction step, assume that the hypothesis is true for some i . Let the $(i+1)$ st edge to be considered be $e = [u, v]$. Two cases arise.

Case 1: u and v belong to the same component of M_i . Then e forms a cycle with M_i , and hence it is not included in M_{i+1} . Since u and v belong to the same component of M_i by the hypothesis they also belong to the same corresponding component of G_i . If e is not included in G_{i+1} , nothing has happened, and the hypothesis remains true. If e is included in G_{i+1} , it gets added to the same component of G_{i+1} and the hypothesis remains true.

Case 2: u and v belong to different components of M_i . Then e does not form a cycle with M_i , and hence it is included in M_{i+1} . In this case two components of M_i merge to form one component of M_{i+1} . Since u and v belong to different components of M_i by the hypothesis they also belong to different corresponding components of G_i . In other words, the distance from u to v in G_i is infinite. Thus e will clearly pass the test in the SPANNER algorithm and be included in G_{i+1} . In this case the two components of G_i merge to form one component of G_{i+1} . Thus the hypothesis remains true, and the inductive proof is complete.

Finally, since $M_{\text{Size}(G)} = MST(G)$, and $G_{\text{Size}(G)} = G$, the lemma is proved. \square

In proving the two theorems, Lemma 2 is used in proving size sparseness, while Lemmas 3 and 4 are used in proving weight sparseness of the resulting subgraphs.

We still need two more lemmas. Let the size of a face of a planar graph be the number of edges encountered while traversing around the boundary of the face. We are allowed to repeat an edge along the boundary, in case the face contains several biconnected components, connected by edges. Lemma 5 bounds the size of a planar graph, given a minimum face size.

Lemma 5. If all the faces of an n -vertex connected planar graph G have sizes $\geq r$, then $\text{Size}(G) \leq (n-2)(1+2/r-2)$.

Proof. Intuitively, the lemma shows that the minimum face size acts as a parameter for size sparseness of planar graphs. To prove this, let m be the size of the graph. Euler's formula for planar graphs states that $n - m + f = 2$, where f is the number of faces in the graph. Thus $f = m + 2 - n$. If we traverse the boundary of each face and mark the edges encountered, every edge in the graph will eventually be marked twice. This is because edges that have different faces adjacent on both sides are counted twice, the traversal of each adjacent face contributing once. Edges with the same face on both sides are also counted twice during the traversal of the same face. Since the size of each face is $\geq r$, we have $f \cdot r \leq 2m$. Substituting for f from above, we get

$$(m + 2 - n)r \leq 2m$$
$$m \leq (n - 1) \left(1 + \frac{2}{r - 2} \right)$$

Thus the lemma is proved. We observe that the lemma has meaning only for $r \leq 2n - 2$, because this is the maximum possible face size of a connected planar graph. This happens when the graph is a tree. \square

Our next lemma is from extremal graph theory, and is easily derivable from Theorem 3.7, Chapter III, in [Bo]. Define the girth of a graph as the size (number of edges) of its smallest simple cycle. This lemma provides an upper bound on the size of a (not necessarily planar) graph with a given girth.

Lemma 6. Let G be an n -vertex graph with girth $> r$. Then $\text{Size}(G) < n^{2n/r - 2}$.

Thus the girth acts as a parameter for size sparseness of general graphs. At this stage we are ready to prove our two theorems. We first prove Theorem 2, then use some of the ideas involved to prove Theorem 1.

Proof of Theorem 2. We run the SPANNER algorithm on a given n -vertex planar graph G and a $t > 0$, after setting $r = 2t + 1$. By Lemma 1, the resulting graph has a stretch factor at most $2t + 1$. Also, by Lemma 2, the girth of the output is

$F.e. r=4$
 $\Rightarrow n \cdot \sqrt{24-2} = n^2$
 $(n) > 0$ Pffs (M) (d)
 $r=9$

[M] and Lubotzky et al. [LPS]. Although they use very similar constructions, the result stated in [M] is more general and provides the required improvement. We now briefly describe the graphs constructed in [M].

Let p be an odd prime. Let $m > 2\sqrt{p}$ be an even integer relatively prime to p . The graph $G^{p,m}$ obtained from p and m is a $(p+1)$ -regular graph on $n \leq m^3/8$ vertices. The girth of $G^{p,m}$ is $g \geq r = 4 \log_p(m/2)$ [M]. Thus, $G^{p,m}$ is an n -vertex graph with more than $\frac{1}{2}n^{1+4/3r}$ edges.

Using these graphs, we improve the bound given in Lemma 8:

Lemma 9. For every pair of integers $r, n \geq 3$, there exists an n -vertex graph G with $girth\ g \geq r$ and with more than $\frac{1}{8}n^{1+4/3r}$ edges.

Proof. If $r = 3$ we can take the complete bipartite graph on n vertices as G .

If $r = 4$ we can take the complete bipartite graph with partition of sizes $\lfloor n/2 \rfloor$ and $\lceil n/2 \rceil$ as G .

From now on we consider $r \geq 5$.

Case 1: $n < (3^{r/4} + 2)^3$. Then

$$\frac{1}{8}n^{1+4/3r} < \frac{1}{8}n(3^{r/4} + 2)^{4r} < \frac{1}{8}n \cdot 5 < n - 1.$$

Thus, we can take a tree on n vertices as G .

Case 2: $n \geq (3^{r/4} + 2)^3$. By Bertrand's postulate there exists a prime p such that $\frac{1}{2}(n^{1/3} - 2)^{4r} \geq p \leq (n^{1/3} - 2)^{4r}$. Since $\frac{1}{2}(n^{1/3} - 2)^{4r} \geq \frac{1}{2}((3^{r/4} + 2) - 2)^{4r} > \frac{1}{2}$, p is an odd prime.

Let m be the least even integer such that $m \geq 2p^{r/4}$ and m is relatively prime to p . Clearly, $m \leq 2p^{r/4} + 4$.

Consider the graph $G^{p,m}$ as described before. This graph is a $(p+1)$ -regular graph on $n' \leq m^3/8$ vertices and it has girth $g \geq \log_p(m/2)$.

Let G be the graph consisting of the union of $\lfloor n/n' \rfloor$ disjoint copies of $G^{p,m}$ and $n \bmod n'$ new isolated vertices. Since $n' \geq m^3/8$ we have that

$$n' \leq \frac{(2p^{r/4} + 4)^3}{8} \leq \frac{(2(n^{1/3} - 2) + 4)^3}{8} = n.$$

This implies that $n \bmod n' < n/2$.

Thus, we obtain that

$$Size(G) \geq \frac{(n - (n \bmod n'))(p + 1)}{2} > \frac{n}{4} (p + 1) \geq \frac{n}{4} \frac{(n^{1/3} - 2)^{4r} + 2}{2} > \frac{n}{4} \frac{n^{4r/3}}{2} = \frac{1}{8}n^{1+4/3r},$$

Handwritten notes: $n \leq m^3/8$, $n' \leq m^3/8$, $n \bmod n' < n/2$, $n' \leq n$, $n' \geq m^3/8$, $n' \leq (2p^{r/4} + 4)^3/8$, $n' \leq (2(n^{1/3} - 2) + 4)^3/8 = n$, $n \bmod n' < n/2$, $Size(G) \geq \frac{(n - (n \bmod n'))(p + 1)}{2} > \frac{n}{4} (p + 1) \geq \frac{n}{4} \frac{(n^{1/3} - 2)^{4r} + 2}{2} > \frac{n}{4} \frac{n^{4r/3}}{2} = \frac{1}{8}n^{1+4/3r}$

which proves the bound on the size. It remains to show that G has girth as promised.

The girth g of G is equal to the girth of $G^{p,m}$. Thus

$$g \geq 4 \log_p \left(\frac{m}{2} \right) \geq 4 \log_p \left(\frac{2p^{r/4}}{2} \right) \geq r.$$

If we put unit weights on the edges of these graphs, then the weight of the minimum spanning tree becomes $n - 1$. Thus, the following is immediate from Lemmas 7 and 9.

Theorem 4. For every pair of integers $t \geq 0$ and $n \geq 3$, there exists an n -vertex graph G for which every $(2t+1)$ -spanner G' is such that

- (1) $Size(G') > \frac{1}{8}n^{1+4/3(2t+3)}$,
- (2) $Weight(G') > \frac{1}{8}Weight(MST(G))n^{4/3(2t+3)}$.

This theorem shows that the size result of Theorem 1 is tight up to a constant factor in the exponent of n , while, for the weight, the gap between lower and upper bounds is larger.

Our next result concerns planar graphs with arbitrary positive weights.

Theorem 5. For every integer $t \geq 1$, there are infinitely many values of n , for which there exist n -vertex planar graphs G with unit edge weights, such that every $(2t+1)$ -spanner G' satisfies

- (1) $Size(G') = \Omega(n(1 + 1/t))$,
- (2) $Weight(G') = \Omega(Weight(MST(G))(1 + 1/t))$.

Proof. Let n take the values $k(t+1) + 2$ for every $k \geq 2$. For every admissible pair of t and n , we construct an n -vertex planar graph G as follows. Lay out two special vertices, u and v , on the plane. Connect them via k chains of $t+1$ vertices each. Let all edges have unit weights.

Clearly, the graph is planar. Furthermore, the girth is $2t+4$. Now consider any proper subgraph of G , say G' . Let $[a, b]$ be an edge in $G - G'$. Because the girth is $\geq 2t+4$, any alternate path from a to b in G' has length at least $2t+3$. This means that there can be no proper spanner of G with a stretch factor $2t+1$. To achieve this stretch factor or less, we have to include the whole graph.

We now prove the size bound of G . By construction, G consists of $(n-2)/(t+1)$ edge-disjoint chains. Since each chain contains $t+2$ edges, the total number of edges in G is $(t+2)(n-2)/(t+1)$.

Thus the size of the graph is $\Omega(n(1 + 1/t))$. The weight bound is trivially true, since unit edge weights imply that the weight of the minimum spanning tree is $n-1$. This theorem shows that the size and weight results of Theorem 2 are tight. \square

Handwritten notes: $t=1$, $(3-2)n$, $\Rightarrow n^{1+4/3(2+3)}$

4. Steiner Spanners: A Lower Bound

In what follows we prove lower bound results for various generalizations of spanners. Consider graphs with arbitrary positive weights. Let $Distance(u, v, G)$ be the distance from vertex u to vertex v in graph G . Let $G_1 = (V_1, E_1)$ and $G_2 = (V_2, E_2)$ be two graphs with V_1 a subset of V_2 . G_2 is called a Steiner t -spanner of G_1 if, for all $u, v \in V_1$, $Distance(u, v, G_2) \leq t \cdot Distance(u, v, G_1)$. Thus, a Steiner spanner is not simply a subgraph of G_1 , rather it may contain auxiliary vertices and edges. An important point that is the Steiner spanner is not allowed to "cheat," that is, paths in G_2 are never shorter than those in the original graph G_1 , though they may use auxiliary vertices and edges.

For special graphs Steiner spanners can be substantially smaller than simple spanners. For example, consider the complete n -vertex graph with unit edge weights. Clearly, every simple 1-spanner requires all edges, but the star graph with one auxiliary vertex attached to all n original vertices via n additional edges of weight $\frac{1}{2}$ is a Steiner 1-spanner with only n edges. However, the following lower bound result shows that such constructions are not always possible for all graphs. In fact, there exist graphs such that Steiner spanners cannot be much smaller than simple spanners.

Theorem 6. For every pair of integers $t \geq 1$ and $n \geq 3$, there exists an n -vertex graph G with unit edge weights for which every Steiner t -spanner requires $\Omega(n/\log n) \cdot n^{1+4/3t+2}$ edges.

Before we prove the theorem, we need some definitions and lemmas. Let $t \geq 1$ and let m be a positive integer. Let g be a function that maps each unweighted n -vertex graph to a string of m bits. We say g is an (n, t, m) -compressor if the following is true. For all pairs of graphs G_1, G_2 , if $g(G_1) = g(G_2)$, then, for all pairs of vertices u, v , $Distance(u, v, G_1) \leq t \cdot Distance(u, v, G_2)$. Informally, the compressor is just like a hash function. It partitions all graphs into groups, such that in each group the graphs have approximately the same distances between any given pair of vertices. Clearly, as t increases, it should be possible to construct compressors with fewer groups. The following lemma provides a lower bound on the number of groups, given any t .

Lemma 10. For every pair of integers $t \geq 1$ and $n \geq 3$, (n, t, m) -compressors exist only for $m \geq \frac{1}{t} \cdot n^{1+4/3t+2}$.

Proof. By Lemma 9, for every pair of integers $t \geq 1$ and $n \geq 3$, there exists an n -vertex graph with girth $\geq t + 2$ and at least $\frac{1}{t} \cdot n^{1+4/3t+2}$ edges. Let $G = (V, E)$ be such a graph, and let G_1, G_2 be two different subgraphs of G . By the girth condition of G , both subgraphs have to belong to different groups of the compressor. Thus the total number of groups is $\geq 2|E|$, thus $m \geq |E|$. \square

Two further observations are useful in proving the theorem. First, existence of spanners (even Steiner spanners) for all graphs implies the existence of a com-

pressor. For example, consider a collection of t -spanners, one for each n -vertex graph. Let the set of spanners (without repetitions) be $\{G_1, \dots, G_k\}$. Then, setting $g(G)$ to be the binary representation of the integer i if G_i is a spanner of G , yields an $(n, t, \lceil \log k \rceil)$ -compressor.

Second, we encode an unweighted graph as a bit string in the following way. Every n -vertex graph with m edges can be encoded by $2 \cdot \lceil \log n \rceil \cdot m$ bits, by representing every edge as a pair of $\lceil \log n \rceil$ bit strings, each in turn representing a vertex.

Proof of Theorem 6. We first prove the theorem for unweighted spanners, then prove it for weighted spanners. To prove the first part, consider any collection of t -spanners, one for each n -vertex graph. Since we are interested in sparse Steiner spanners of n -vertex graphs, even with auxiliary vertices these spanners should have no more than $n + 2 \cdot n(n - 1)/2 = n^2$ vertices. Let m be the size of the largest spanner in this set. Each spanner can be encoded in $\leq 4 \cdot \lceil \log n \rceil \cdot m$ bits. Thus by our previous observation, these spanners imply the existence of an $(n, t, 4 \cdot \lceil \log n \rceil \cdot m)$ -compressor, but, by Lemma 10, $4 \cdot \lceil \log n \rceil \cdot m \geq \frac{1}{t} \cdot n^{1+4/3t+2}$. Solving for m proves the theorem.

To prove the second part, again consider a collection of t -spanners, one for each n -vertex graph with unit edge weights. As before these spanners cannot have more than n^2 vertices. However, since the edges have weights, we have to devise a different encoding scheme. Informally the idea is to encode the weights themselves as bit strings. The problem arises because weights may require an arbitrary number of bits. However, we show that by ignoring very large weights, and rounding off the remainder in a certain way, we can get suitably encodable spanners with slightly larger stretch factors.

Consider any t -spanner G' in the set. First notice that any edge with weight larger than $n - 1$ can be removed without affecting the stretch factor. Let G'' be another spanner derived from G' in the following way. For each edge e replace its weight by $Weight''(e) = \lceil n^3 \cdot Weight(e) \rceil / n^3$. Notice that $Weight''(e) < Weight(e) + 1/n^3$.

We now claim that G'' is a $(t + 1/n)$ -spanner. This is easy to see because any path in G'' has at most $n^2 - 1$ edges, and each weight in G'' is at most $1/n^3$ greater than the corresponding weight in G' .

G'' is more convenient to code than G' . Let m be the size of the largest G'' in the set. Each edge weight is a rational, where the numerator is an integer between 1 and $n^4 - n^3$ and the denominator is n^3 . Thus each weight can be encoded by less than $\lceil \log n^4 \rceil \leq 4 \lceil \log n \rceil$ bits. Thus G'' can be encoded by at most $8 \cdot \lceil \log n \rceil \cdot m$ bits. So these spanners imply the existence of an $(n, t + 8 \cdot \lceil \log n \rceil \cdot m)$ -compressor. Applying Lemma 10 and solving for m , the theorem follows. \square

Note that we have no lower bounds on the weight sparseness of Steiner spanners. We now generalize spanners in another direction. Let $t \geq 1$. $G' = (V, E)$ is a t -approximator of $G = (V, E)$ if, for all $u, v \in V$, $1/t \cdot Distance(u, v, G) \leq Distance(u, v, G') \leq t \cdot Distance(u, v, G)$. The following theorem provides a lower bound on the size of approximators.

Theorem 10. For every $t > 1$, dimension d , and norm $\|\cdot\|$ of \mathbb{R}^d , there exists a constant $c(t, d, \|\cdot\|)$ such that every finite set V in \mathbb{R}^d has a t -spanner with at most $c \cdot n$ edges.

In the rest of this section we estimate the angle $\delta(\varepsilon)$ and thus prove Lemma 11. The following lemmas will be useful in the proof.

Lemma 12. Let $\alpha, \beta, \gamma > 0$, $\alpha + \beta + \gamma = \pi$, $\gamma \leq \beta$, and $\alpha \leq 2\pi/3$. Then

$$\frac{\sin \beta - \sin \alpha}{\sin \gamma} \geq 1 - \alpha.$$

Proof. Fix $\alpha > 0$ and define

$$g(\gamma) = \frac{\sin(\pi - \alpha - \gamma) - \sin \alpha}{\sin \gamma}.$$

We have to show that $g(\gamma) \geq 2\alpha$ for all γ in $(0, (\pi - \alpha)/2]$. By the rule of Hôpital,

$$\lim_{\gamma \rightarrow 0} g(\gamma) = \lim_{\gamma \rightarrow 0} \frac{-\cos(\pi - \alpha - \gamma)}{\cos \gamma} = -\cos(\pi - \alpha) = \cos \alpha = 1 - 2 \sin^2 \frac{\alpha}{2} \geq 1 - \alpha.$$

For the other extreme in the interval we get

$$\begin{aligned} g\left(\frac{\pi - \alpha}{2}\right) &= \frac{\sin(\pi - \alpha)/2 - \sin \alpha}{\sin(\alpha - \alpha)/2} = 1 - \frac{\sin \alpha}{\sin(\pi - \alpha)/2} \\ &= 1 - \frac{\sin \alpha}{\cos \alpha/2} = 1 - 2 \sin \frac{\alpha}{2} \geq 1 - \alpha. \end{aligned}$$

We complete the proof by showing that the derivative $g'(\gamma)$ has no zero in the open interval $(0, (\pi - \alpha)/2)$. Clearly,

$$g'(\gamma) = \frac{-\cos(\pi - \alpha - \gamma) \sin \gamma - [\sin(\pi - \alpha - \gamma) - \sin \alpha] \cos \gamma}{\sin^2 \gamma}.$$

An application of the addition formula for the sinus yields

$$g'(\gamma) = \frac{-\sin(\pi - \alpha) + \sin \alpha \cos \gamma}{\sin^2 \gamma}.$$

The numerator equals $-\sin \alpha + \sin \alpha \cos \gamma$ and is smaller than 0 for all $0 < \gamma < (\pi - \alpha)/2$. \square

Lemma 13. In the Euclidean norm $\|\cdot\|_2$, every triangle with side lengths $c \leq b$ and angle $\alpha < 2\pi/3$ satisfies $(b - a)/c \geq 1 - \alpha$.

Proof. Apply the sinus theorem $a/\sin \alpha = b/\sin \beta = c/\sin \gamma$ to Lemma 12. \square

In the rest of the section we use the following notation. The angle between any two points u and v with respect to the origin in \mathbb{R}^d will be written as $\angle(u, v)$. Also $\sin \angle(u, v)$ will be written in short as $\sin(u, v)$.

Lemma 14. Consider the norm $\|\cdot\|_2$ and any other norm $\|\cdot\|$ in \mathbb{R}^d , such that $\|u\| \leq q\|u\|_2$ for all u in \mathbb{R}^d , where $q > 0$ is an appropriate constant. Then

$$\left| \frac{\|u\|}{\|u\|_2} - \frac{\|v\|}{\|v\|_2} \right| \leq \sqrt{2}q|\sin(u, v)| \quad \text{for all } u, v \text{ other than the origin.}$$

Proof. The triangle inequalities are $\|u\| - \|v\| \leq \|u - v\|$ and $\|u\| - \|v\| \leq \|u + v\|$. Without loss of generality assume $\|u\|_2 = \|v\|_2 = 1$ and $\|u\| \geq \|v\|$. Let $\langle \cdot, \cdot \rangle$ denote the standard scalar product.

Case 1: $\langle u, v \rangle \geq 0$. By the first triangle inequality we get

$$\begin{aligned} \left| \frac{\|u\|}{\|u\|_2} - \frac{\|v\|}{\|v\|_2} \right|^2 &\leq \|u - v\|^2 \leq q^2\|u - v\|_2^2 = q^2\langle u - v, u - v \rangle \\ &= q^2(\langle u, u \rangle + \langle v, v \rangle - 2\langle u, v \rangle) = 2q^2(1 - \langle u, v \rangle). \end{aligned}$$

On the other hand,

$$\begin{aligned} |\sin(u, v)|^2 &= 1 - \cos^2(u, v) = 1 - \langle u, v \rangle^2 = (1 - \langle u, v \rangle)(1 + \langle u, v \rangle) \\ &\geq 1 - \langle u, v \rangle. \end{aligned}$$

Case 2: $\langle u, v \rangle < 0$. This case is solved analogously by applying the other triangle inequality. \square

Lemma 15. Let $\|u\|_2 \leq \|u\| \leq q\|u\|_2$ for some $q \geq 1$ and all u in \mathbb{R}^d . Let $0 < \varepsilon < 1/(1 + 2\sqrt{2})$. Then

$$\frac{\|u\| - \|u - v\|}{\|v\|} > 1 - (1 + 2\sqrt{2})\varepsilon$$

for all u, v other than the origin with $\|u\|_2 \geq \|v\|_2$ and $\angle(u, v) < \varepsilon/q$.

Proof. By Lemma 13 we have

$$\frac{\|u\|_2 - \|u - v\|_2}{\|v\|_2} > 1 - \frac{\varepsilon}{q}. \tag{1}$$

Consider the expression $(\|u\| - \|v - u\|)/\|v\|$. By Lemma 14 and multiplying the numerator and denominator by $\|u\|_2/\|u\|$ we get

$$\begin{aligned} \frac{\|u\| - \|v - u\|}{\|v\|} &\geq \frac{\|u\| - \|v - u\|_2 [\|u\|/\|u\|_2 + \sqrt{2q} \sin(u, v - u)]}{\|v\|_2 [\|u\|/\|u\|_2 + \sqrt{2q} \sin(u, v)]} \\ &= \frac{\|u\|_2 - \|v - u\|_2 [1 + (\|u\|_2/\|u\|) \sqrt{2q} \sin(u, v - u)]}{\|v\|_2 [1 + (\|u\|_2/\|u\|) \sqrt{2q} \sin(u, v)]} \\ &\geq \frac{\|u\|_2 - \|v - u\|_2 [1 + \sqrt{2q} \sin(u, v - u)]}{\|v\|_2 [1 + \sqrt{2q} \sin(u, v)]} \\ &= \frac{1 + \sqrt{2q} |\sin(u, v)|}{1} \\ &\quad \times \left[\frac{\|u\|_2 - \|v - u\|_2}{\|v\|_2} - \frac{\|v - u\|_2 \sqrt{2q} |\sin(u, v - u)|}{\|v\|_2} \right]. \end{aligned} \tag{2}$$

By the sinus theorem we have $\sin(u, v - u) = (\|v\|_2/\|v - u\|_2) \sin(u, v)$. Together with (1) this shows that (2) is greater than

$$\begin{aligned} \frac{1}{1 + \sqrt{2q} \sin(\epsilon/q)} \cdot \left[1 - \frac{\epsilon}{q} - \sqrt{2q} \sin\left(\frac{\epsilon}{q}\right) \right] &\geq \frac{1}{1 + \sqrt{2q}(\epsilon/q)} \cdot \left[1 - \epsilon - \sqrt{2q} \frac{\epsilon}{q} \right] \\ &= \frac{1 - (1 + \sqrt{2})\epsilon}{1 + \sqrt{2}\epsilon} \\ &> 1 - (1 + 2\sqrt{2})\epsilon. \end{aligned} \quad \square$$

Proof of Lemma 11. Let the norm under consideration be $\|\cdot\|$ such that $\|u\|_2 \leq \|u\| \leq q\|u\|_2$ for some $q \geq 1$ and all u in R^d . Select the angle $\delta(\epsilon)$ to be ϵ/q . Recall the manner in which a path from v to u is constructed in the graph G . Let the vertices along this path be $v = z_0, \dots, z_p = u$. The following are true for all i :

$$\angle(z_p - z_i, z_{i+1} - z_i) < \frac{\epsilon}{q} \quad \text{and} \quad \|z_p - z_i\|_2 \geq \|z_{i+1} - z_i\|_2.$$

By Lemma 15 we get

$$\|z_{i+1} - z_i\| < \frac{1}{1 - (1 + 2\sqrt{2})\epsilon} [\|z_p - z_i\| - \|z_p - z_{i+1}\|].$$

Thus,

$$\sum_{i=0}^{p-1} \|z_{i+1} - z_i\| < \frac{1}{1 - (1 + 2\sqrt{2})\epsilon} \sum_{i=0}^{p-1} [\|z_p - z_i\| - \|z_p - z_{i+1}\|]$$

A direct application of Lemma 13 shows that in the case of the $\|\cdot\|_2$ norm we get $1/(1 - \epsilon)$ spanners with $s(\delta(\epsilon)) \cdot n$ edges. In R^2 , $s(\delta(\epsilon))$ is known to be $\lfloor 2\pi/\epsilon \rfloor + 1$.

$$\begin{aligned} &= \frac{1}{1 - (1 + 2\sqrt{2})\epsilon} [\|z_p - z_0\| - \|z_p - z_p\|] \\ &= \frac{1}{1 - (1 + 2\sqrt{2})\epsilon} \|z_p - z_0\|. \end{aligned} \quad \square$$

6. Open Problems

We conclude with some open problems.

(1) In Theorem 1 the bound in the weight does not agree with the lower bound as nicely as in the other results. Can it be improved? We feel that our strategy of dividing the graph into planar components cannot be extended to yield the optimal answer.

(2) Even in the other results, there are gaps between upper and lower bounds. For instance, in Theorem 1 the upper bound for the size is $O(n^{1+1/n})$, while the lower bound is $\Omega(n^{1+4/(e+2)})$. Only for $t = 1$ are the known bounds, $\Theta(n^{3/2})$, of the same order. The lower bound was proved in [LJ], and the upper bound is mentioned in [PUP].

(3) For dimensions higher than 2, Euclidean spanners with linear sizes exist. Do Euclidean spanners exist with weights within a constant multiple of the weight of the minimum spanning tree?

(4) What spanners do random graphs have? The Euclidean random case has been examined in [SV].

(5) Consider R^d , some fixed norm, and $t > 1$. What are the worst (or at least bad) point configurations V with respect to the number of edges in optimal t -spanners?

(6) Do spanners have other applications?

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**Independent Collections of Translates of Boxes and a
Lecture due to Grünbaum***

Daniel Q. Naiman¹ and Henry P. Wynn²

¹Department of Mathematical Sciences, The Johns Hopkins University,
Baltimore, MD 21218, USA
dqn@ms.jhu.edu

²City University, London, England
h.p.wynn@city.ac.uk

Abstract. A collection of n sets A_1, \dots, A_n is said to be *independent* provided every subcollection of the form $X_1 \cap \dots \cap X_n$ is nonempty, where each X_i is either A_i or A_i^c . We give a simple characterization for when translates of a given box form an independent set in \mathbb{R}^d . We use this to show that the largest number of such boxes forming an independent set in \mathbb{R}^d is given by $\lfloor 3d/2 \rfloor$ for $d \geq 2$. This settles in the negative a conjecture of Grünbaum (1975), which states that the maximum size of an independent collection of sets homothetic to a fixed convex set C in \mathbb{R}^d is $d + 1$. It also shows that the bound of 2d of Rényi *et al.* (1951) for the maximum number of boxes that necessarily translates of a given one) with sides parallel to the coordinate axes form an independent collection in \mathbb{R}^d can be improved for these special collections.

Introduction

A finite collection of sets $\mathcal{A} = \{A_i, i = 1, \dots, n\}$ is said to be *independent* if for every subcollection $X_i \neq \emptyset$, whenever $X_i = A_i$ or A_i^c for $i = 1, \dots, n$. Marczewski (1947) describes how the notion of independence is related to the problem of extending measures. Indeed, \mathcal{A} is independent if and only if the every subcollection of $P(A_i) \in [0, 1]$, $i = 1, \dots, n$, there exists a unique probability measure P defined on the algebra of sets generated by \mathcal{A} under which A_1, \dots, A_n are probabilistically independent.

Use of independence leads to simplifications to the classical inclusion-exclusion formula. In fact, Naiman and Wynn (1991b) show for any collection of sets that the maximum size of an independent subcollection is k , then the indicator

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