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► **To cite this version:**

Georgia Avarikioti, Ioannis Z. Emiris, Loukas Kavouras, Ioannis Psarros. High-dimensional approximate r-nets. SODA: ACM/SIAM Symposium on Discrete Algorithms, Jan 2017, Barcelone, Spain. hal-01636792

HAL Id: hal-01636792

<https://hal.inria.fr/hal-01636792>

Submitted on 17 Oct 2018

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High-dimensional approximate r -nets

Georgia Avarikioti* Ioannis Z. Emiris† Loukas Kavouras‡ Ioannis Psarros§

May 9, 2017

Abstract

The construction of r -nets offers a powerful tool in computational and metric geometry. We focus on high-dimensional spaces and present a new randomized algorithm which efficiently computes approximate r -nets with respect to Euclidean distance. For any fixed $\epsilon > 0$, the approximation factor is $1 + \epsilon$ and the complexity is polynomial in the dimension and subquadratic in the number of points. The algorithm succeeds with high probability. More specifically, the best previously known LSH-based construction of Eppstein et al. [EHS15] is improved in terms of complexity by reducing the dependence on ϵ , provided that ϵ is sufficiently small. Our method does not require LSH but, instead, follows Valiant's [Val15] approach in designing a sequence of reductions of our problem to other problems in different spaces, under Euclidean distance or inner product, for which r -nets are computed efficiently and the error can be controlled. Our result immediately implies efficient solutions to a number of geometric problems in high dimension, such as finding the $(1 + \epsilon)$ -approximate k th nearest neighbor distance in time subquadratic in the size of the input.

Keywords: Metric geometry, High dimension, Approximation algorithms, r -nets, Locality-sensitive hashing

1 Introduction

We study r -nets, a powerful tool in computational and metric geometry, with several applications in approximation algorithms. An r -net for a metric space $(X, \|\cdot\|)$, $|X| = n$ and for numerical parameter r is a subset $R \subseteq X$ such that the closed $r/2$ -balls centered at the points of R are disjoint, and the closed r -balls around the same points cover all of X . We define approximate r -nets analogously. Formally,

Definition 1. *Given a pointset $X \subseteq \mathbb{R}^d$, a distance parameter $r \in \mathbb{R}$ and an approximation parameter $\epsilon > 0$, a $(1 + \epsilon)r$ -net of X is a subset $R \subseteq X$ s.t. the following properties hold:*

1. (packing) *For every $p, q \in R$, $p \neq q$, we have that $\|p - q\|_2 \geq r$.*
2. (covering) *For every $p \in X$, there exists a $q \in R$ s.t. $\|p - q\|_2 \leq (1 + \epsilon)r$.*

Previous Work. Finding r -nets can be addressed naively by considering all points of X unmarked and, while there remains an unmarked point p , the algorithm adds it to R and marks all other points within distance r from p . The performance of this algorithm can be improved by using grids and hashing [Har04]. However, their complexity remains too large when dealing with big data in high dimension. The naive algorithm is quadratic in n and the grid approach is in

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$O(d^{d/2}n)$, hence it is relevant only for constant dimension d [HR15]. In [HP05], they show that an approximate net hierarchy for an arbitrary finite metric can be computed in $O(2^{ddim}n \log n)$, where $ddim$ is the doubling dimension. This is satisfactory when doubling dimension is constant, but requires a vast amount of resources when it is high.

When the dimension is high, there is need for algorithms with time complexity polynomial in d and subquadratic in n . One approach, which computes $(1 + \epsilon)r$ -nets in high dimension is that of [EHS15], which uses the Locality Sensitive Hashing (LSH) method of [AI08]. The resulting time complexity is $\tilde{O}(dn^{2-\Theta(\epsilon)})$, where $\epsilon > 0$ is quite small and \tilde{O} hides polylogarithmic factors.

In general, high dimensional analogues of classical geometric problems have been mainly addressed by LSH. For instance, the approximate closest pair problem can be trivially solved by performing n approximate nearest neighbor (ANN) queries. For sufficiently small ϵ , this costs $\tilde{O}(dn^{2-\Theta(\epsilon)})$ time, due to the complexity factor of an LSH query. Several other problems have been reduced to ANN queries [GIV01]. Recently, Valiant [Val12], [Val15] presented an algorithm for the approximate closest pair problem in time $\tilde{O}(dn^{2-\Theta(\sqrt{\epsilon})})$. This is a different approach in the sense that while LSH exploits dimension reduction through random projections, the algorithm of [Val15] is inspired by high dimensional phenomena. One main step of the algorithm is that of projecting the pointset up to a higher dimension.

Our Contribution. We present a new randomized algorithm that computes approximate r -nets in time subquadratic in n and polynomial in the dimension, and improves upon the complexity of the best known algorithm. Our method does not employ LSH and, with probability $1 - o(1)$, it returns $R \subset X$, which is a $(1 + \epsilon)r$ -net of X .

We reduce the problem of an approximate r -net for arbitrary vectors (points) under Euclidean distance to the same problem for vectors on the unit sphere. Then, depending on the magnitude of distance r , an algorithm handling “small” distances or an algorithm handling “large” distances is called. These algorithms reduce the Euclidean problem of r -nets on unit vectors to finding an r -net for unit vectors under inner product (Section 3). This step requires that the multiplicative $1 + \epsilon$ approximation of the distance corresponds to an additive $c\epsilon$ approximation of the inner product, for suitable constant $c > 0$.

Next, we convert the vectors having unit norm into vectors with entries $\{-1, +1\}$ (Section 2). This transformation is necessary in order to apply the Chebyshev embedding of [Val15], an embedding that damps the magnitude of the inner product of “far” vectors, while preserving the magnitude of the inner product of “close” vectors. For the final step of the algorithm, we first apply a procedure that allows us to efficiently compute $(1 + \epsilon)$ -nets in the case where the number of “small” distances is large. Then, we apply a modified version of the **Vector Aggregation** algorithm of [Val15], that exploits fast matrix multiplication, so as to achieve the desired running time.

In short, we extend Valiant’s framework [Val15] and we compute r -nets in time $\tilde{O}(dn^{2-\Theta(\sqrt{\epsilon})})$, thus improving on the exponent of the LSH-based construction [EHS15], when ϵ is small enough. This improvement by $\sqrt{\epsilon}$ in the exponent is the same as the complexity improvement obtained in [Val15] over the LSH-based algorithm for the approximate closest pair problem.

Our study is motivated by the observation that computing efficiently an r -net leads to efficient solutions for several geometric problems, specifically in approximation algorithms. In particular, our extension of r -nets in high dimensional Euclidean space can be plugged in the framework of [HR15]. The new framework has many applications, notably the k th nearest neighbor distance problem, which we solve in $\tilde{O}(dn^{2-\Theta(\sqrt{\epsilon})})$.

Paper Organization. Section 2 presents an algorithm for computing an approximate net with respect to the inner product for a set of unit vectors. Section 3 translates the problem of finding r -nets under Euclidean distance to the same problem under inner product. In Section

4, we discuss applications of our construction and possible future work. Omitted proofs are included in the Appendices.

We use $\|\cdot\|$ to denote the Euclidean norm $\|\cdot\|_2$ throughout the paper.

2 Points on a sphere under inner product

In this section, we design an algorithm for constructing an approximate ρ -net of vectors on the sphere under inner product. To that end, we reduce the problem to constructing an approximate net under absolute inner product for vectors that lie on the vertices of a unit hypercube.

Since our ultimate goal is a solution to computing r -nets with respect to Euclidean distance, we allow additive error in the approximation, which under certain assumptions, translates to multiplicative error in Euclidean distance. In the following, we define rigorously the notion of approximate ρ -nets under inner product.

Definition 2. For any $X \subset \mathbb{S}^{d-1}$, an approximate ρ -net for $(X, \langle \cdot, \cdot \rangle)$, with additive approximation parameter $\epsilon > 0$, is a subset $C \subseteq X$ which satisfies the following properties:

- for any two $p \neq q \in C$, $\langle p, q \rangle < \rho$, and
- for any $x \in X$, there exists $p \in C$ s.t. $\langle x, p \rangle \geq \rho - \epsilon$.

One relevant notion is that of ϵ -kernels [AHPV05]. In ϵ -kernels, one is interested in finding a subset of the input pointset, which approximates its directional width. Such constructions have been extensively studied when the dimension is low, due to their relatively small size.

2.1 Crude approximate nets

In this subsection we develop our basic tool, which is based on the Vector Aggregation Algorithm by [Val15]. This tool aims to compute approximate ρ -nets with multiplicative error, as opposed to what we have set as our final goal for this section, namely to bound additive error. Moreover, in the context of this subsection, two vectors are close to each other when the magnitude of their inner product is large, and two vectors are far from each other when the magnitude of their inner product is small. Let $|\langle \cdot, \cdot \rangle|$ denote the magnitude of the inner product of two vectors.

Definition 3. For any $X = [x_1, \dots, x_n], X' = [x'_1, \dots, x'_n] \subset \mathbb{R}^{d \times n}$, a crude approximate ρ -net for $(X, X', |\langle \cdot, \cdot \rangle|)$, with multiplicative approximation factor $c > 1$, is a subset $C \subseteq [n]$ which satisfies the following properties:

- for any two $i \neq j \in C$, $|\langle x_i, x'_j \rangle| < c\rho$, and
- for any $i \in [n]$, there exists $j \in C$ s.t. $|\langle x_i, x'_j \rangle| \geq \rho$.

Vector Aggregation follows the exposition of [Val15]. The main difference is that, instead of the “compressed” matrix $Z^T Z$, we use the form $X^T Z$, where Z derives from vector aggregation. Both forms encode the information in the Gram matrix $X^T X$. The matrix $X^T Z$ is better suited for our purposes, since each row corresponds to an input vector instead of an aggregated subset; this extra information may be useful in further problems.

Vector Aggregation

Input: $X = [x_1, \dots, x_n] \in \mathbb{R}^{d \times n}$, $X' = [x'_1, \dots, x'_n] \in \mathbb{R}^{d \times n}$, $\alpha \in (0, 1)$, $\tau > 0$.

Output: $n \times n^{1-\alpha}$ matrix W and random partition $S_1, \dots, S_{n^{1-\alpha}}$ of $\{x_1, \dots, x_n\}$.

- Randomly partition $[n]$ into $n^{1-\alpha}$ disjoint subsets, each of size n^α , denoting the sets $S_1, \dots, S_{n^{1-\alpha}}$.
- For each $i = 1, 2, \dots, 78 \log n$:
 - Select n coefficients $q_1, \dots, q_n \in \{-1, +1\}$ at random.
 - Form the $d \times n^{1-\alpha}$ matrix Z^i with entries $z_{j,k}^i = \sum_{l \in S_k} q_l \cdot x'_{j,l}$
 - $W^i = X^T Z^i$
- Define the $n \times n^{1-\alpha}$ matrix W with $w_{i,j} = \text{quartile}(|w_{i,j}^1|, \dots, |w_{i,j}^{78 \log n}|)$.
- Output W and $S_1, \dots, S_{n^{1-\alpha}}$.

Theorem 4. Let $X \in \mathbb{R}^{d \times n}$, $X' \in \mathbb{R}^{d \times n}$, $\alpha \in (0, 1)$, $\tau > 0$ the input of **Vector Aggregation**. Then, the algorithm returns a matrix W of size $n \times n^{1-\alpha}$ and a random partition $S_1, \dots, S_{n^{1-\alpha}}$, which with probability $1 - O(1/n^3)$ satisfies the following:

- For all $j \in [n]$ and $k \in [n^{1-\alpha}]$, if $\forall u \in S_k$, $|\langle x_j, u \rangle| \leq \tau$ then $|w_{j,k}| < 3 \cdot n^\alpha \tau$.
- For all $j \in [n]$ and $k \in [n^{1-\alpha}]$ if $\exists u \in S_k$, $|\langle x_j, u \rangle| \geq 3n^\alpha \tau$ then $|w_{j,k}| \geq 3 \cdot n^\alpha \tau$.

Moreover, the algorithm runs in time $\tilde{O}(dn + n^{2-\alpha} + \text{MatrixMul}(n \times d, d \times n^{1-\alpha}))$.

For the case of pointsets with many “small” distances, we rely crucially on the fact that the expected number of near neighbors for a randomly chosen point is large. So, if we iteratively choose random points and delete these and their neighbors, we will end up with a pointset which satisfies the property of having sufficiently few “small” distances. Then, we apply **Vector Aggregation**.

Crude ApprxNet

Input: $X = [x_1, \dots, x_n] \in \mathbb{R}^{d \times n}$, $X' = [x'_1, \dots, x'_n] \in \mathbb{R}^{d \times n}$, $\alpha \in (0, 1)$, $\tau > 0$.

Output: $C' \subseteq [n]$, $F' \subseteq [n]$.

- $C \leftarrow \emptyset$, $F_1 \leftarrow \emptyset$, $F_2 \leftarrow \{x_1, \dots, x_n\}$
- Repeat $n^{0.5}$ times:
 - Choose a column x_i uniformly at random.
 - $C \leftarrow C \cup \{x_i\}$.
 - Delete column i from matrix X and column i from matrix X' .
 - Delete each column k from matrix X , X' s.t. $|\langle x_i, x'_k \rangle| \geq \tau$.
 - If there is no column k from matrix X s.t. $|\langle x_i, x'_k \rangle| \geq \tau$, then $F_1 \leftarrow F_1 \cup \{x_i\}$
- Run **Vector Aggregation** with input X , X' , α , τ and output W , $S_1, \dots, S_{n^{1-\alpha}}$.
- For each of the remaining columns $i = 1, \dots$:
 - For any $|w_{i,j}| \geq 3n^\alpha \tau$:
 - * If more than $n^{1.7}$ times in here, output “ERROR”.

- * Compute inner products between x_i and vectors in S_j . For each vector $x'_k \in S_j$ s.t. $x'_k \neq x_i$ and $|\langle x_i, x'_k \rangle| \geq \tau$, delete row k and $F_2 \leftarrow F_2 \setminus \{x_i\}$.
- $C \leftarrow C \cup \{x_i\}$
- Output indices of C and $F \leftarrow \{F_1 \cup F_2\}$.

Theorem 5. *On input $X = [x_1, \dots, x_n] \in \mathbb{R}^{d \times n}$, $X' = [x'_1, \dots, x'_n] \in \mathbb{R}^{d \times n}$, $\alpha \in (0, 1)$, $\tau > 0$, **Crude ApprxNet**, computes a crude $3n^\alpha$ -approximate τ -net for X, X' , following the notation of Definition 3. The algorithm costs time:*

$$\tilde{O}(n^{2-\alpha} + d \cdot n^{1.7+\alpha} + \text{MatrixMul}(n \times d, d \times n^{1-\alpha})),$$

and succeeds with probability $1 - O(1/n^{0.2})$. Additionally, it outputs a set $F \subseteq R$ with the following property: $\{x_i \mid \forall x_j \neq x_i \ |\langle x_j, x_i \rangle| < \tau\} \subseteq F \subseteq \{x_i \mid \forall x_j \neq x_i \ |\langle x_j, x_i \rangle| < n^\alpha \tau\}$.

Proof. We perform $n^{0.5}$ iterations and for each, we compare the inner products between the randomly chosen vector and all other vectors. Hence, the time needed is $O(dn^{1.5})$.

In the following, we denote by X_i the number of vectors which have “large” magnitude of the inner product with the randomly chosen point in the i th iteration. Towards proving correctness, suppose first that $\mathbb{E}[X_i] > 2n^{0.5}$ for all $i = 1, \dots, n^{0.5}$. The expected number of vectors we delete in each iteration of the algorithm is more than $2n^{0.5} + 1$. So, after $n^{0.5}$ iterations, the expected total number of deleted vectors will be greater than n . This means that if the hypothesis holds for all iterations we will end up with a proper net.

Now suppose that there is an iteration j where $\mathbb{E}[X_j] \leq 2n^{0.5}$. After all iterations, the number of “small” distances are at most $n^{1.5}$ on expectation. By Markov’s inequality, when the **Vector Aggregation** algorithm is called, the following is satisfied with probability $1 - n^{-0.2}$:

$$|\{(i, k) \mid |\langle x_i, x'_k \rangle| \geq \tau, i \neq k\}| \leq n^{1.7}.$$

By Theorem 4 and the above discussion, the number of entries in the matrix W that we need to visit is at most $n^{1.7}$. For each entry, we perform a brute force which costs dn^α .

Now notice that the first iteration stores centers c and deletes all points p for which $|\langle c, p \rangle| \geq \tau$. Hence, any two centers c, c' satisfy $|\langle c, p \rangle| < \tau$. In the second iteration, over the columns of W , notice that by Theorem 4, for any two centers c, c' we have $|\langle c, c' \rangle| < 3n^\alpha \tau$. \square

2.2 Approximate inner product nets

In this subsection, we show that the problem of computing ρ -nets for the inner product of unit vectors reduces to the less natural problem of Definition 3, which refers to the magnitude of the inner product.

The first step consists of mapping the unit vectors to vectors in $\{-1, 1\}^{d'}$. The mapping is essentially Charikar’s LSH scheme [Cha02]. Then, we apply the Chebyshev embedding of [Val15] in order to achieve gap amplification, and finally we call algorithm **Crude ApprxNet**, which will now return a proper ρ -net with additive error.

Theorem 6 ([Val15]). *There exists an algorithm with the following properties. Let $d' = O(\frac{\log n}{\delta^2})$ and $Y \in \mathbb{R}^{d' \times n}$ denote its output on input X , δ , where X is a matrix whose columns have unit norm, with probability $1 - o(1/n^2)$, for all pairs $i, j \in [n]$, $\left| \langle Y_i, Y_j \rangle / d' - \left(1 - 2 \cdot \cos^{-1}(\langle X_i, X_j \rangle) / \pi \right) \right| \leq \delta$, where X_i, Y_i denote the i th column of X and Y respectively. Additionally, the runtime of the algorithm is $O(\frac{dn \log n}{\delta^2})$.*

The following theorem provides a randomized embedding that damps the magnitude of the inner product of “far” vectors, while preserving the magnitude of the inner product of “close” vectors. The statement is almost verbatim that of [Val15, Prop.6] except that we additionally establish an asymptotically better probability of success. The proof is the same, but since we claim stronger guarantees on success probability, we include the complete proof in Appendix B.

Theorem 7. *Let Y, Y' be the matrices output by algorithm “Chebyshev Embedding” on input $X, X' \in \{-1, 1\}^{d \times n}$, $\tau^+ \in [-1, 1]$, $\tau^- \in [-1, 1]$ with $\tau^- < \tau^+$, integers q, d' . With probability $1 - o(1/n)$ over the randomness in the construction of Y, Y' , for all $i, j \in [n]$, $\langle Y_i, Y'_j \rangle$ is within $\sqrt{d'} \log n$ from the value $T_q \left(\frac{\langle X_i, X'_j \rangle / d' - \tau^-}{\tau^+ - \tau^-} 2 - 1 \right) \cdot d' \cdot (\tau^+ - \tau^-)^q / 2^{3q-1}$, where T_q is the degree- q Chebyshev polynomial of the first kind. The algorithm runs in time $O(d' \cdot n \cdot q)$.*

Inner product ApprxNet

Input: $X = [x_1, \dots, x_n]$ with each $x_i \in \mathbb{S}^{d-1}$, $\rho \in [-1, 1]$, $\epsilon \in (0, 1/2]$.

Output: Sets $C, F \subseteq [n]$.

- If $\rho \leq \epsilon$, then:
 - $C \leftarrow \emptyset, F \leftarrow \emptyset, W \leftarrow \{x_1, \dots, x_n\}$
 - While $W \neq \emptyset$:
 - * Choose arbitrary vector $x \in W$.
 - * $W \leftarrow W \setminus \{y \in W \mid \langle x, y \rangle \geq \rho - \epsilon\}$
 - * $C \leftarrow C \cup \{x\}$
 - * If $\forall y \in W, \langle x, y \rangle < \rho - \epsilon$ then $F \leftarrow F \cup \{x\}$
 - Return indices of C, F .
- Apply Theorem 6 for input X , $\delta = \epsilon/2\pi$ and output $Y \in \{-1, 1\}^{d' \times n}$ for $d' = O(\log n / \delta^2)$.
- Apply Theorem 7 for input Y , $d'' = n^{0.2}$, $q = 50^{-1} \log n$, $\tau^- = -1$, $\tau^+ = 1 - \frac{2 \cos^{-1}(\rho - \epsilon)}{\pi} + \delta$ and output Z, Z' .
- Run algorithm Crude ApprxNet with input $\tau = 3n^{0.16}$, $\alpha = \sqrt{\epsilon}/500$, Z, Z' and output C, F .
- Return C, F .

Theorem 8. *The algorithm Inner product ApprxNet, on input $X = [x_1, \dots, x_n]$ with each $x_i \in \mathbb{S}^{d-1}$, $\rho \in [-1, 1]$ and $\epsilon \in (0, 1/2]$, computes an approximate ρ -net with additive error ϵ , using the notation of Definition 2. The algorithm runs in time $\tilde{O}(dn + n^{2-\sqrt{\epsilon}/600})$ and succeeds with probability $1 - O(1/n^{0.2})$. Additionally, it computes a set F with the following property: $\{x_i \mid \forall x_j \neq x_i \langle x_j, x_i \rangle < \rho - \epsilon\} \subseteq F \subseteq \{x_i \mid \forall x_j \neq x_i \langle x_j, x_i \rangle < \rho\}$.*

3 Approximate nets in high dimensions

In this section, we translate the problem of computing r -nets in $(\mathbb{R}^d, \|\cdot\|)$ to the problem of computing ρ -nets for unit vectors under inner product. One intermediate step is that of computing r -nets for unit vectors under Euclidean distance.

3.1 From arbitrary to unit vectors

In this subsection, we show that if one is interested in finding an r -net for $(\mathbb{R}^d, \|\cdot\|)$, it is sufficient to solve the problem for points on the unit sphere. One analogous statement is used in [Val15], where they prove that one can apply a randomized mapping from the general Euclidean space to points on a unit sphere, while preserving the ratio of distances for any two pairs of points. The claim derives by the simple observation that an r -net in the initial space can be approximated by computing an $\epsilon r/c$ -net on the sphere, where c is the maximum norm of any given point envisaged as vector. Our exposition is even simpler since we can directly employ the analogous theorem from [Val15].

Corollary 9. *There exists an algorithm, `Standardize`, which, on input a $d \times n$ matrix X with entries $x_{i,j} \in \mathbb{R}$, a constant $\epsilon \in (0, 1)$ and a distance parameter $r \in \mathbb{R}$, outputs a $m' \times n$ matrix Y , with columns having unit norm and $m' = \log^3 n$, and a distance parameter $\rho \in \mathbb{R}$, such that a ρ -net of Y is an approximate r -net of X , with probability $1 - o(1/\text{poly}(n))$.*

3.2 Approximate nets under Euclidean distance

In this subsection, we show that one can translate the problem of computing an r -net for points on the unit sphere under Euclidean distance, to finding an r -net for unit vectors under inner product as defined in Section 2. Moreover, we identify the subset of the r -net which contains the centers that are approximately far from any other point. Formally,

Definition 10. *Given a set of points X and $\epsilon > 0$, a set $F \subseteq X$ of $(1 + \epsilon)$ -approximate r -far points is defined by the following property: $\{x \in X \mid \forall y \in X \ \|x - y\| > (1 + \epsilon)r\} \subseteq F \subseteq \{x \in X \mid \forall y \in X \ \|x - y\| > r\}$.*

If r is greater than some constant, the problem can be immediately solved by the law of cosines. If r cannot be considered as constant, we distinguish cases $r \geq 1/n^{0.9}$ and $r < 1/n^{0.9}$. The first case is solved by a simple modification of an analogous algorithm in [Val15, p.13:28]. The second case is not straightforward and requires partitioning the pointset in a manner which allows computing r -nets for each part separately. Each part has bounded diameter which implies that we need to solve a “large r ” subproblem.

Theorem 11. *There exists an algorithm, `ApprxNet(Large radius)`, which, for any constant $\epsilon \in (0, 1/2]$, $X \subset \mathbb{S}^{d-1}$ s.t. $|X| = n$, outputs a $(1 + \epsilon)r$ -net and a set of $(1 + \epsilon)$ -approximate r -far points with probability $1 - O(1/n^{0.2})$. Additionally, provided $r > 1/n^{0.9}$ the runtime of the algorithm is $\tilde{O}(dn^{2-\Theta(\sqrt{\epsilon})})$.*

Let us now present an algorithm which translates the problem of finding an r -net for $r < 1/n^{0.9}$ to the problem of computing an r -net for $r \geq 1/n^{0.9}$. The main idea is that we compute disjoint subsets S_i , which are far enough from each other, so that we can compute r -nets for each S_i independently. We show that for each S_i we can compute $T_i \subseteq S_i$ which has bounded diameter and $T'_i \subseteq S_i$ such that T_i, T'_i are disjoint, each point in T_i is far from each point in T'_i , and $|T'_i| \leq 3|S_i|/4$. It is then easy to find r -nets for T_i by employing the `ApprxNet(Large radius)` algorithm. Then, we recurse on T'_i which contains a constant fraction of points from $|S_i|$. Then, we cover points in $S_i \setminus (T_i \cup T'_i)$ and points which do not belong to any S_i .

ApprxNet(Small radius)

Input: $X = [x_1, \dots, x_n]^T$ with each $x_i \in \mathbb{S}^{d-1}$, $r < 1/n^{0.9}$, $\epsilon \in (0, 1/2]$.

Output: Sets $R, F \subseteq [n]$.

1. Project points on a uniform random unit vector and consider projections p_1, \dots, p_n which wlog correspond to $x_1, \dots, x_n \in \mathbb{R}^d$.
2. Traverse the list as follows:
 - If $|\{j \mid p_j \in [p_i - r, p_i]\}| \leq n^{0.6}$ or $i = n$:
 - If $|\{j \mid p_j < p_i\}| \leq n^{0.9}$ remove from the list all points p_j s.t. $p_j < p_i - r$ and save set $K = \{x_j \mid p_j \in [p_i - r, p_i]\}$.
 - If $|\{j \mid p_j < p_i\}| > n^{0.9}$ save sets $K_i = \{x_j \mid p_j \in [p_i - r, p_i]\} \cup K$, $S_i = \{x_j \mid p_j < p_i - r\} \setminus K$ and remove projections of S_i and K_i from the list.
3. After traversing the list if we have not saved any S_i go to 5; otherwise for each S_i :
 - For each $u \in S_i$, sample $n^{0.1}$ distances between u and randomly chosen $x_k \in S_i$. Stop if for the selected $u \in S_i$, more than $1/3$ of the sampled points are in distance $\leq rn^{0.6}$. This means that one has found u s.t. $|\{x_k \in S_i, \|u - x_k\| \leq rn^{0.6}\}| \geq |S_i|/4$ with high probability. If no such point was found, output "ERROR".
 - Let $0 \leq d_1 \leq \dots \leq d_{|S_i|}$ be the distances between u and all other points in S_i . Find $c \in [rn^{0.6}, 2rn^{0.6}]$ s.t. $|\{j \in [n] \mid d_j \in [c, c + r]\}| < n^{0.4}$, store $W_i = \{x_j \mid d_j \in [c, c + r]\}$, and remove W_i from S_i .
 - Construct the sets $T_i = \{x_j \in S_i \mid d_j < c\}$ and $T'_i = \{x_j \in S_i \mid d_j > c + r\}$.
 - For T_i , subtract u from all vectors in T_i , run **Standardize**, then **ApprxNet (Large radius)**, both with $\epsilon/4$. Save points which correspond to output at R_i, F_i respectively.
 - Recurse on T'_i the whole algorithm, and notice that $|T'_i| \leq 3|S_i|/4$. Save output at R'_i, F'_i respectively.
4. Let $R \leftarrow \bigcup_i R_i \cup R'_i$ and $F \leftarrow \bigcup_i F_i \cup F'_i$. Return to the list p_1, \dots, p_n .
 - (a) Remove from F all points which cover at least one point from $\bigcup_i W_i$ or $\bigcup_i K_i$.
 - (b) Delete all points $(\bigcup_i T_i) \setminus (\bigcup_i R_i)$, and $(\bigcup_i T'_i) \setminus (\bigcup_i R'_i)$.
 - (c) For each i delete all points in W_i covered by R_i , or covered by R'_i .
 - (d) For each i delete all points in K_i covered by R .
 - (e) Finally delete R from the list. Store the remaining points at F' .
5. $R' \leftarrow \emptyset$. Traverse the list as follows: For each p_i , check the distances from all x_j s.t. $p_j \in [p_i - r, p_i]$.
 - If $\exists x_j \in R' : \|x_i - x_j\| \leq r$, delete x_i from the list, set $F' \leftarrow F' \setminus \{x_i, x_j\}$ and continue traversing the list.
 - If there is no such point x_j then $R \leftarrow R \cup \{x_i\}$ and continue traversing the list.
6. Output indices of $R \leftarrow R \cup R'$ and $F \leftarrow F \cup F'$.

Theorem 12. For any constant $\epsilon > 0$, $X \subset \mathbb{S}^{d-1}$ s.t. $|X| = n$, and $r < 1/n^{0.9}$, **ApprxNet (Small radius)** will output a $(1 + \epsilon)r$ -net and a set of $(1 + \epsilon)$ -approximate r -far points in time $\tilde{O}(dn^{2-\Theta(\sqrt{\epsilon})})$, with probability $1 - o(1/n^{0.04})$.

Proof. Note that points in S_i had projections p_i in sets of contiguous intervals of width r ; each interval had $\geq n^{0.6}$ points, hence the diameter of the projection of S_i is $\leq n^{0.4}r$. By the Johnson

Lindenstrauss Lemma [Das03] we have that for $v \in \mathbb{S}^{d-1}$ chosen uniformly at random:

$$\Pr \left[\langle u, v \rangle^2 \leq \frac{\|u\|^2}{n^{0.4}} \right] \leq \frac{\sqrt{d}\sqrt{e}}{n^{0.2}}.$$

Hence, $\mathbb{E}[\{\|x_k - x_j\| \geq n^{0.6}r \text{ and } \|p_k - p_j\| \leq n^{0.4}r\}] \leq |S_i|^2 \cdot \frac{\sqrt{ed}}{n^{0.2}}$, and the probability

$$\Pr[\{\|x_k - x_j\| \geq n^{0.6}r \text{ and } \|p_k - p_j\| \leq n^{0.4}r\} \geq |S_i|^{1.95}] \leq |S_i|^{0.05} \cdot \frac{\sqrt{ed}}{n^{0.2}} \leq \frac{\sqrt{ed}}{n^{0.15}}.$$

Taking a union bound over all sets S_i yields a probability of failure $o(1/n^{0.045})$. This implies that (for large enough n , which implies large enough $|S_i|$) at least

$$\binom{|S_i|}{2} - |S_i|^{1.95} \geq \frac{|S_i|^2}{4}$$

distances between points in S_i are indeed small ($\leq n^{0.6}r$). Hence, there exists some point $p_k \in S_i$ which $(n^{0.6}r)$ -covers $|S_i|/2$ points. For each possible p_k we sample $n^{0.1}$ distances to other points, and by Chernoff bounds, if a point $(n^{0.6}r)$ -covers a fraction of more than $1/2$ of the points in S_i , then it covers more than $n^{0.1}/3$ sampled points with high probability. Similarly, if a point $(n^{0.6}r)$ -covers a fraction of less than $1/4$ of the points in S_i , then it covers less than $n^{0.1}/3$ sampled points with high probability. More precisely, for some fixed $u \in S_i$, let $X_j = 1$ when for the j th randomly chosen point $v \in S_i$, it holds $\|u - v\| \leq n^{0.6}r$ and let $X_j = 0$ otherwise. Then, for $Y = \sum_{j=1}^{n^{0.1}} X_j$, it holds:

$$\mathbb{E}[Y] \geq n^{0.1}/2 \implies \Pr[Y \leq n^{0.1}/3] \leq \exp(-\Theta(n^{0.1})),$$

$$\mathbb{E}[Y] \leq n^{0.1}/4 \implies \Pr[Y \geq n^{0.1}/3] \leq \exp(-\Theta(n^{0.1})).$$

Since for any point $x \in T_i$ and any point $y \in T'_i$ we have $\|x - y\| > r$, the packing property of r -nets is preserved when we build r -nets for T_i and T'_i independently. For each T_i , we succeed in building r -nets with probability $1 - O(1/n^{0.2})$. By a union bound over all sets T_i , we have a probability of failure $O(1/n^{0.1})$. Furthermore, points which belong to sets W_i and K_i are possibly covered and need to be checked.

For the analysis of the runtime of the algorithm, notice that step 4b costs time $O(d \cdot (\sum_i |T_i| + \sum_i |T'_i|)) = O(dn)$. Then, step 4c costs time $O(d \cdot \sum_i |W_i| \cdot |T_i| + d \cdot \sum_i |W_i| \cdot |T'_i|) = O(dn^{1.4})$. Finally, notice that we have at most $n^{0.1}$ sets K_i . Each K_i contains at most $2n^{0.6}$ points, hence checking each point in $\bigcup_i K_i$ with each point in R costs $O(dn^{1.7})$.

Now regarding step 5, consider any interval $[p_i - r, p_i]$ in the initial list, where all points are projected. If $|\{j \mid p_j \in [p_i - r, p_i]\}| \leq 2n^{0.9}$ then the i th iteration in step 5 will obviously cost $O(n^{0.9})$, since previous steps only delete points. If $|\{j \mid p_j \in [p_i - r, p_i]\}| > 2n^{0.9}$, we claim that $|\{j < i \mid p_j \in [p_i - r, p_i] \text{ and } K_j \text{ is created}\}| \leq 1$. Consider the smallest $j < i$ s.t. K_j is created and $p_j \in [p_i - r, p_i]$. This means that all points p_k , for $k \leq j$, are deleted when p_j is visited. Now assume that there exists integer $l \in (j, i)$ s.t. K_l is created. This means that the remaining points in the interval $[p_l - r, p_l]$ are $\leq n^{0.6}$ and all of the remaining points $p_k < p_l$ are more than $n^{0.9}$. This leads to contradiction, since by the deletion in the j th iteration, we know that all of the remaining points $p_k < p_l$ lie in the interval $[p_l - r, p_l]$.

Now, assume that there exists one $j < i$ s.t. $p_j \in [p_i - r, p_i]$ and K_j is created. Then, when p_i is visited, there at least $2n^{0.9} - n^{0.6} > n^{0.9}$ remaining points in the interval $[p_i - r, p_i]$. Hence, there exists $l \geq i$ for which the remaining points in the interval $[p_i - r, p_i]$ are contained in $S_l \cup K_l$. Hence in this case, in step 5, there exist at most $O(n^{0.6})$ points which are not deleted and belong to the interval $[p_i - r, p_i]$. Now assume that there does not exist any $j < i$ s.t.

$p_j \in [p_i - r, p_i]$ and K_j is created. This directly implies that there exists $l \geq i$ for which the remaining points in the interval $[p_i - r, p_i]$ are contained in $S_l \cup K_l$.

At last, the total time of the above algorithm is dominated by the calls to the construction of the partial r -nets of the sets T_i . Thus, the total running time is $O(\sum_i |T_i|^{2-\Theta(\sqrt{\epsilon})} + \sum_i |T_i|^{2-\Theta(\sqrt{\epsilon})}) = O(\sum_i |T_i|^{2-\Theta(\sqrt{\epsilon})} + \sum_i (3|T_i|/4)^{2-\Theta(\sqrt{\epsilon})}) = \tilde{O}(n^{2-\Theta(\sqrt{\epsilon})})$. Finally, taking a union bound over all recursive calls of the algorithm we obtain a probability of failure $o(1/n^{0.04})$. \square

We now present an algorithm for an $(1 + \epsilon)r$ -net for points in \mathbb{R}^d under Euclidean distance.

ApprxNet

Input: Matrix $X = [x_1, \dots, x_n]$ with each $x_i \in \mathbb{R}^d$, parameter $r \in \mathbb{R}$, constant $\epsilon \in (0, 1/2]$.

Output: $R \subseteq \{x_1, \dots, x_n\}$

- Let Y, r' be the output of algorithm **Standardize** on input X, r with parameter $\epsilon/4$.
- If $r \geq 1/n^{0.9}$ run **ApprxNet(Large radius)** on input $Y, \epsilon/4, r'$ and return points which correspond to the set R .
- If $r < 1/n^{0.9}$ run **ApprxNet(Small radius)** on input $Y, \epsilon/4, r'$ and return points which correspond to the set R .

Theorem 13. *Given n points in \mathbb{R}^d , a distance parameter $r \in \mathbb{R}$ and an approximation parameter $\epsilon \in (0, 1/2]$, with probability $1 - o(1/n^{0.04})$, **ApprxNet** will return a $(1 + \epsilon)r$ -net, R , in $\tilde{O}(dn^{2-\Theta(\sqrt{\epsilon})})$ time.*

Proof. The theorem is a direct implication of Theorems 11, 12, 22. \square

Theorem 14. *Given $X \subset \mathbb{R}^d$ such that $|X| = n$, a distance parameter $r \in \mathbb{R}$ and an approximation parameter $\epsilon \in (0, 1/2]$, there exists an algorithm, **DelFar**, that will return, with probability $1 - o(1/n^{0.04})$, a set F' with the following properties in $\tilde{O}(dn^{2-\Theta(\sqrt{\epsilon})})$ time:*

- If for a point $p \in X$ it holds that $\forall q \neq p, q \in X$ we have $\|p - q\| > (1 + \epsilon)r$, then $p \notin F'$.
- If for a point $p \in X$ it holds that $\exists q \neq p, q \in X$ s.t. $\|p - q\| \leq r$, then $p \in F'$.

4 Applications and Future work

Concerning applications, in [HR15], they design an approximation scheme, which solves various distance optimization problems. The technique employs a grid-based construction of r -nets which is linear in n , but exponential in d . The main prerequisite of the method is the existence of a linear-time decider (formally defined in Appendix G). The framework is especially interesting when the dimension is constant, since the whole algorithm costs time linear in n which, for some problems, improves upon previously known near-linear algorithms. When the dimension is high, we aim for polynomial dependency on d , and subquadratic dependency on n .

Let us focus on the problem of approximating the k th nearest neighbor distance.

Definition 15. *Let $X \subset \mathbb{R}^d$ be a set of n points, approximation error $\epsilon > 0$, and let $d_1 \leq \dots \leq d_n$ be the nearest neighbor distances. The problem of computing an $(1 + \epsilon)$ -approximation to the k th nearest neighbor distance asks for a pair $x, y \in X$ such that $\|x - y\| \in [(1 - \epsilon)d_k, (1 + \epsilon)d_k]$.*

Now we present an approximate decider for the problem above. This procedure combined with the framework we mentioned earlier, which employs our net construction, results in an efficient solution for this problem in high dimension.

kth NND Decider

Input: $X \subseteq \mathbb{R}^d$, constant $\epsilon \in (0, 1/2]$, integer $k > 0$.

Output: An interval for the optimal value $f(X, k)$.

- Call $\text{DelFar}(X, \frac{r}{1+\epsilon/4}, \epsilon/4)$ and store its output in W_1 .
- Call $\text{DelFar}(X, r, \epsilon/4)$ and store its output in W_2 .
- Do one of the following:
 - If $|W_1| > k$, then output “ $f(X, k) < r$ ”.
 - If $|W_2| < k$, then output “ $f(X, k) > r$ ”.
 - If $|W_1| \leq k$ and $|W_2| \geq k$, then output “ $f(X, k) \in [\frac{r}{1+\epsilon/4}, \frac{1+\epsilon/4}{r}]$ ”.

Theorem 16. *Given a pointset $X \subseteq \mathbb{R}^d$, one can compute a $(1 + \epsilon)$ -approximation to the k -th nearest neighbor in $\tilde{O}(dn^{2-\Theta(\sqrt{\epsilon})})$, with probability $1 - o(1)$.*

To the best of our knowledge, this is the first high dimensional solution for this problem. Setting $k = n$ and applying Theorem 16 one can compute the *farthest nearest neighbor* in $\tilde{O}(dn^{2-\Theta(\sqrt{\epsilon})})$ with high probability.

Concerning future work, let us start with the problem of finding a greedy permutation. A permutation $\Pi = \langle \pi_1, \pi_2, \dots \rangle$ of the vertices of a metric space $(X, \|\cdot\|)$ is a *greedy permutation* if each vertex π_i is the farthest in X from the set $\Pi_{i-1} = \langle \pi_1, \dots, \pi_{i-1} \rangle$ of preceding vertices. The computation of r -nets is closely related to that of the greedy permutation.

The k -center clustering problem asks the following: given a set $X \subseteq \mathbb{R}^d$ and an integer k , find the smallest radius r such that X is contained within k balls of radius r . By [EHS15], a simple modification of our net construction implies an algorithm for the $(1 + \epsilon)$ approximate greedy permutation in time $\tilde{O}(dn^{2-\Theta(\sqrt{\epsilon})} \log \Phi)$ where Φ denotes the spread of the pointset. Then, approximating the greedy permutation implies a $(2 + \epsilon)$ approximation algorithm for k -center clustering problem. We expect that one can avoid any dependencies on Φ .

Acknowledgment.

I.Z. Emiris acknowledges partial support by the EU H2020 research and innovation programme, under the Marie Skłodowska-Curie grant agreement No 675789: Network “ARCADES”.

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A Proof of Theorem 4

Lemma 17 (Anti-concentration). *Let $q_1, \dots, q_t \in \{-1, 1\}$ be chosen independently and uniformly at random, and let $a_1, \dots, a_t \in \mathbb{R}$ s.t. $|a_1| = \max_i |a_i|$. Then,*

$$\Pr\left[\left|\sum_{i=1}^t q_i \cdot a_i\right| \geq |a_1|\right] \geq 1/2.$$

Proof. Consider a given assignment for q_2, \dots, q_t . Then if

$$\sum_{i=2}^t q_i \cdot a_i = 0 \implies \left|\sum_{i=1}^t q_i \cdot a_i\right| = |q_1 \cdot a_1| = |a_1|.$$

Otherwise,

$$\Pr\left[\left|\sum_{i=1}^t q_i \cdot a_i\right| \geq |a_1|\right] \geq \Pr[\text{sign}(q_1 \cdot a_1) = \text{sign}(\sum_{i=2}^t q_i \cdot a_i)] = 1/2.$$

□

Proof of Theorem 4. Notice that

$$w_{j,k}^i = \sum_{x_i \in S_k} q_i \cdot \langle x_j, x_i \rangle$$

and since $q_1, \dots, q_{|S_k|} \in \{-1, 1\}$ are independent and chosen uniformly at random, we obtain

$$\mathbb{E}[w_{j,k}^i] = 0.$$

If $\forall u \in S_k, |\langle x_j, u \rangle| \leq \tau$, then

$$\text{Var}(w_{j,k}^i) = \mathbb{E}[(w_{j,k}^i)^2] \leq n^{2\alpha} \tau^2$$

By Chebyshev's inequality:

$$\Pr[|w_{j,k}^i| \geq 3 \cdot n^\alpha \tau] \leq 1/9$$

With m repetitions, the number of successes N , that is the number of indices i for which $|w_{j,k}^i| \leq 3 \cdot n^\alpha \tau$, follows the binomial distribution. Hence,

$$\Pr[N \leq 3m/4] \leq \exp(-m/26).$$

We consider as bad event the event that for some j, k , more than 25% of the repetitions fail, that is $|w_{j,k}^i| \geq 3 \cdot n^\alpha \tau$. By the union bound, this probability is $\leq n^{2-\alpha} \cdot \exp(-m/26)$, which for $m \geq 78 \log n$ implies a probability of failure $\leq 1/n^3$.

Now consider x_j , and $x_l \in S_k$ s.t. $|\langle x_j, x_l \rangle| \geq 3 \cdot n^\alpha \tau$, then by Lemma 17, with probability 1/2, $|w_{j,k}^i| \geq 3 \cdot n^\alpha \tau$. We consider as bad event the event that for j, l , more than 75% of the repetitions fail, that is $|w_{j,k}^i| \leq 3 \cdot n^\alpha \tau$. Hence,

$$\Pr[N \leq m/4] \leq \exp(-m/8),$$

which for $m \geq 78 \log n$ implies a probability of failure $\leq 1/n^3$.

The runtime of the algorithm is dominated, up to polylogarithmic factors, by the computation of matrix Z , taking time $O(dn)$, the computation of matrix W , taking time n^{2-a} , or the computation of the product W^i , taking time $MatrixMul(n \times d, d \times n^{1-a})$.

□

B Proof of Theorem 7

We refer to [Val15, Algorithm 3: Chebyshev Embedding]. The proof is the same with that of [Val15], apart from indicating that the probability of success is actually $1 - o(1/n)$ instead of $1 - o(1)$ as stated in [Val15]. While $1 - o(1/n)$ probability of success is enough for our purposes, even better probability bounds can be achieved.

The fact that all inner products are concentrated within $\pm\sqrt{m} \log n$ about their expectations follows from the fact that each row of Y, Y' is generated identically and independently from the other rows, and all entries of these matrices are ± 1 ; thus, each inner product is a sum of independent and identically distributed random ± 1 random variables, and we can apply the basic Chernoff bound to each inner product, and then a union bound over the $O(n^2)$ inner products. Let $X_i \in \pm 1$ i.i.d. random variables. The basic Chernoff bound gives probability,

$$\Pr\left[\left|\sum_{i=1}^{m'} X_i - \mathbb{E}\left[\sum_{i=1}^{m'} X_i\right]\right| > \sqrt{m'} \log n\right] \leq 2 \cdot \exp(-\Theta(\log^2 n)) = o(1/n^3).$$

Given this concentration, we now analyze the expectation of the inner products. Let u, u' be columns of X, X' , respectively, and v, v' the corresponding columns of Y, Y' . Letting $x = \langle u, u' \rangle / m$, we argue that by [Val15, Lemma 3.3], $\mathbb{E}[v, v'] = m' \sum_{i=1}^q \frac{x - c_i}{2}$ (1), where c_i is the location of the i th root of the q th Chebyshev polynomial after the roots have been scaled to lie in the interval $[\tau^-, \tau^+]$. To see why this is the case, note that each coordinate of u, u' is generated by computing the product of q random variables that are all ± 1 ; namely, a given entry of u is given by $\prod_{l=1}^q s_v(l)$, with the corresponding entry of u' given by $\prod_{l=1}^q s_{v'}(l)$. Note that for $i \neq j$, $s_v(i)$ is independent of $s_v(j)$ and $t_{v'}(j)$, although by construction, $s_v(i)$ and $t_{v'}(i)$ are not independent. We now argue that $\mathbb{E}[s_v(i)t_{v'}(i)] = \frac{x - c_i}{2}$, from which Eq. (1) will follow by the fact that the expectation of the product of independent random variables is the product of their expectations.

By construction, in Step (1) of the inner loop of the algorithm, with probability $1/2$, $\mathbb{E}[s_v(i)t_{v'}(i)] = \langle v, v' \rangle / m = x$. Steps (2)(4) ensure that with the remaining $1/2$ probability, $\mathbb{E}[s_v(i)t_{v'}(i)] = \frac{1 - c_i}{2}(1) - \frac{1 + c_i}{2}(-1) = -c_i$. Hence, in aggregate over the randomness of Steps (1)(4), $\mathbb{E}[s_v(i)t_{v'}(i)] = x/2 - c_i/2i$, as claimed, establishing Eq. (1).

To show that Eq. (1) yields the statement of the proposition, we simply reexpress the polynomial $\prod_{i=1}^q \frac{x - c_i}{2}$ in terms of the q th Chebyshev polynomial T_q . Note that the q th Chebyshev polynomial has leading coefficient 2^{q-1} , whereas this expression (as a polynomial in x) has leading coefficient $1/2^q$, disregarding the factor of the dimension m' . If one has two monic degree q polynomials, P and Q where the roots of Q are given by scaling the roots of P by a factor of α , then the values at corresponding locations differ by a multiplicative factor of $1/\alpha^q$; since the roots of T_q lie between $[-1, 1]$ and the roots of the polynomial constructed in the embedding lie between $[\tau^-, \tau^+]$, this corresponds to taking $\alpha = \frac{2}{\tau^+ - \tau^-}$.

C Proof of Theorem 8

Theorem 18 ([Cop97]). *For any positive $\gamma > 0$, provided that $\beta < 0.29$, the product of a $k \times k^\beta$ with a $k^\beta \times k$ matrix can be computed in time $O(k^{2+\gamma})$.*

Corollary 19. *For any positive $\gamma > 0$, provided that $\beta < 0.29 \cdot \alpha < 1$, the product of a $n \times n^\beta$ by a $n^\beta \times n^\alpha$ matrix can be computed in time $O(n^{1+\alpha+\alpha\gamma})$.*

Proof. The idea is to perform $n^{1-\alpha}$ multiplications of matrices of size $n^\alpha \times n^\beta$ and $n^\beta \times n^\alpha$.

Hence, by Theorem 18, the total cost is:

$$O(n^{1-\alpha}(n^{\alpha(2+\gamma)})) = O(n^{1+\alpha+\alpha\gamma}).$$

□

Fact 20. Let $T_q(x)$ denote the q th Chebyshev polynomial of the first kind, then the following hold:

- For $x \in [-1, 1]$, $|T_q(x)| \leq 1$.
- For $\delta \in (0, 1/2]$, $T_q(1 + \delta) \geq \frac{1}{2}e^{q\sqrt{\delta}}$.

Claim 21. For $\rho \in [-1, 1]$, $\epsilon \in (0, 1)$, it holds $\cos^{-1}(\rho - \epsilon) - \cos^{-1}(\rho) \geq \epsilon/2$.

Proof. If $(\rho - \epsilon)^2 \neq 1$ then we have

$$\begin{aligned} \cos^{-1}(\rho - \epsilon) - \cos^{-1}(\rho) &= \int_{\rho-\epsilon}^1 \frac{1}{\sqrt{1-x^2}} dx - \int_{\rho}^1 \frac{1}{\sqrt{1-x^2}} dx = \\ &= \int_{\rho-\epsilon}^{\rho} \frac{1}{\sqrt{1-x^2}} dx = \int_0^{\epsilon} \frac{1}{\sqrt{1-(\rho-\epsilon+y)^2}} dy \geq \int_0^{\epsilon} \frac{1}{\sqrt{1-(\rho-\epsilon)^2}} dy = \frac{\epsilon}{\sqrt{1-(\rho-\epsilon)^2}} \geq \epsilon. \end{aligned}$$

Now if $(\rho - \epsilon)^2 = 1 \implies \rho - \epsilon = -1$ then,

$$\cos^{-1}(\rho - \epsilon) - \cos^{-1}(\rho) = \int_{-1}^{-1+\epsilon} \frac{1}{\sqrt{1-x^2}} dx \geq \frac{\epsilon}{\sqrt{2\epsilon - \epsilon^2}} \geq \epsilon/2.$$

□

Proof of Theorem 8. If $\rho \leq \epsilon$, our approach ensures that for any $x, y \in C$, it holds $\langle x, y \rangle < \rho - \epsilon \leq 0$. We show that $|C| \leq d + 1$, due to a simple packing argument. Let x_1, \dots, x_{d+2} such that $\forall i \neq j \in [d+2]$ we have $\langle x_i, x_j \rangle < 0$. Then, there exist $\lambda_1, \dots, \lambda_{d+1} \in \mathbb{R}$ not all zero for which $\sum_{i=1}^{d+1} \lambda_i x_i = 0$. Now consider two subsets $I, J \subseteq [d+2]$ of indices such that $\forall i \in I, \lambda_i > 0$ and $\forall j \in J, \lambda_j < 0$. We can write $\sum_{i \in I} \lambda_i x_i = \sum_{j \in J} -\lambda_j x_j \implies 0 \leq \langle \sum_{i \in I} \lambda_i x_i, -\sum_{j \in J} \lambda_j x_j \rangle = -\sum_{i \in I, j \in J} \lambda_i \lambda_j \langle x_i, x_j \rangle < 0$ which leads to contradiction. If $J = \emptyset$ (or equivalently if $I = \emptyset$), then $0 = \langle x_{d+2}, \sum_{i \in I} \lambda_i x_i \rangle < 0$, which leads again to contradiction.

We now focus on the case $\rho > \epsilon$. By Theorem 6, with probability $1 - o(1/n^2)$, the matrix Y returned by the corresponding algorithm will have the property that any pair of columns

$$\begin{aligned} \langle X_i, X_j \rangle \geq \rho &\implies \frac{\langle Y_i, Y_j \rangle}{d'} \geq 1 - \frac{2 \cos^{-1}(\rho)}{\pi} - \delta \\ \langle X_i, X_j \rangle \leq \rho - \epsilon &\implies \frac{\langle Y_i, Y_j \rangle}{d'} \leq 1 - \frac{2 \cos^{-1}(\rho - \epsilon)}{\pi} + \delta. \end{aligned}$$

Hence, according to Claim 21, it suffices to set $\delta = \epsilon/3\pi$ in order to distinguish between the two cases:

$$1 - \frac{2 \cos^{-1}(\rho - \epsilon)}{\pi} + 2\delta \leq 1 - \frac{2 \cos^{-1}(\rho)}{\pi} - \delta.$$

Now we set $\tau^+ = 1 - \frac{2 \cos^{-1}(\rho - \epsilon)}{\pi} + \delta > -1$. By Theorem 7, with probability $1 - o(1)$,

$$\langle Y_i, Y_j \rangle \leq \tau^+ d' \leq \implies |\langle Z_i, Z_j \rangle| \leq d'' \frac{2^q}{2^{3q-1}} + \sqrt{d''} \log n \leq 3n^{0.16}$$

for large enough n . Moreover, let Y_i, Y_j s.t. $\langle Y_i, Y_j \rangle \geq (\tau^+ + \delta)d'$. Then,

$$|\langle Z_i, Z'_j \rangle| \geq d'' \cdot T_q\left(1 + 2\frac{\delta}{\tau^+ + 1}\right) \frac{2^q}{2^{3q-1}} - \sqrt{d''} \log n > \frac{1}{2} \cdot T_q\left(1 + 2\frac{\delta}{\tau^+ + 1}\right) \cdot n^{0.16}$$

for large enough n .

Then, by Fact 20,

$$|\langle Z_i, Z'_j \rangle| \cdot n^{-0.16} \geq \frac{1}{4} e^{q\sqrt{\delta}} = \frac{1}{4} n^{\sqrt{\delta}/50} \geq 3n^{\sqrt{\delta}/100} \geq 3n^{\sqrt{\epsilon}/400},$$

where some of the inequalities hold for large enough n .

Now, by Theorems 6, 7, 5 and Corollary 19 the time complexity is $\tilde{O}(dn + n^{2-\sqrt{\epsilon}/600})$, if we set as γ in Corollary 19 a sufficiently small multiple of $\sqrt{\epsilon}$. Finally,, the subroutine with the higher probability of failure is **Crude ApprxNet** and by the union bound, it dominates the total probability of failure. \square

D Proof of Corollary 9

We use an algorithm introduced in [Val15]: its guarantees are stated below.

Theorem 22. [Val15] *There exists an algorithm which on input a $d \times n$ matrix X with entries $x_{i,j} \in \mathbb{R}$ and a constant $\epsilon \in (0, 1)$ outputs a $m' \times n$ matrix Y with columns having unit norm and $m' = \log^3 n$, such that, with probability $1 - o(1/\text{poly}(n))$ for all sets of four columns Y_1, Y_2, Y_3, Y_4 of matrix Y , with X_1, X_2, X_3, X_4 being the corresponding columns of matrix X , it holds that*

$$\frac{\|Y_1 - Y_2\| \|X_3 - X_4\|}{\|Y_3 - Y_4\| \|X_1 - X_2\|} \in [1 - \frac{\epsilon}{10}, 1 + \frac{\epsilon}{10}].$$

Now, let us define two d -dimensional vectors X_{n+1}, X_{n+2} , s.t. $r' = X_{n+1} - X_{n+2}$ and $\|r'\| = r$, where X is a $d \times n$ matrix with entries $x_{i,j} \in \mathbb{R}$ and $r \in \mathbb{R}$ is the radius of the r -net of X . Also, let matrix X' denote the concatenation of X, X_{n+1} and X_{n+2} with size $d \times (n + 2)$. After applying Theorem 22 on input X' and $\epsilon/10$, we define $\rho := \|Y_{n+1} - Y_{n+2}\|$ to be the new radius of Y . Then, we claim that the following hold with probability $1 - o(1/\text{poly}(n))$, which immediately implies Corollary 9:

- For all $X_i, X_j \in X$ and their corresponding $Y_i, Y_j \in Y$, if $\|X_i - X_j\| \leq r$ then $\|Y_i - Y_j\| \leq (1 + \epsilon/10)\rho$.
- For all $X_i, X_j \in X$ and their corresponding $Y_i, Y_j \in Y$, if $\|X_i - X_j\| \geq (1 + \epsilon)r$ then $\|Y_i - Y_j\| \geq (1 + \epsilon/2)\rho$.

Proof of Corollary 9. From Theorem 22, we easily derive that for all $X_i, X_j \in X$ and their corresponding $Y_i, Y_j \in Y$, it holds that

$$\|Y_i - Y_j\| \leq (1 + \epsilon/10) \frac{\|X_i - X_j\|}{r} \rho$$

Therefore, if $\|X_i - X_j\| \leq r$, we have $\|Y_i - Y_j\| \leq (1 + \epsilon/10)\rho$. For the other direction, we use the opposite side of Theorem 22, thus we have that for all $X_i, X_j \in X$ and their corresponding $Y_i, Y_j \in Y$:

$$\|Y_i - Y_j\| \geq (1 - \epsilon/10) \frac{\|X_i - X_j\|}{r} \rho.$$

It follows that $\|X_i - X_j\| \geq (1 + \epsilon)r \Rightarrow \|Y_i - Y_j\| \geq (1 - \epsilon/10)(1 + \epsilon)\rho \Rightarrow \|Y_i - Y_j\| \geq (1 + \epsilon/2)\rho$. \square

E Proof of Theorem 11

ApprxNet (Large radius)

Input: $X = [x_1, \dots, x_n]^T$ with each $x_i \in \mathbb{S}^{d-1}$ with $d = \log^3 n$, $r > 1/n^{0.9}$, $\epsilon \in (0, 1/2)$.
Output: Sets $R, F \subseteq [n]$.

- If $r > 0.2$ run **Inner Product ApprxNet** with error parameter $\epsilon/25$ and $\rho = 1 - \frac{r^2}{2}$.
- Otherwise, define the $d \times n$ matrix Z as follows: for each $i \in [d]$, select $q = \lfloor \frac{\pi}{2 \cos^{-1}(1-r^2/2)} \rfloor$ uniformly random vectors v_1, \dots, v_q and for all $j \in [n]$, set

$$z_{i,j} = \text{sign} \prod_{k=1}^{k=q} X_j^T v_k,$$

where X_j is the j th column of matrix X .

- Run **Inner Product ApprxNet** with $\rho = \left(1 - \frac{2 \cos^{-1}(1-r^2/2)}{\pi}\right)^q$, error parameter $\epsilon/100$ and input matrix Z with all entries scaled by $1/\sqrt{d}$ to make them have unit norm.

Proof of Theorem 11.

In the case of $r > 0.2$ we will show that the $1+\epsilon$ multiplicative approximation on the distance translates to $c\epsilon$ additive approximation to the inner product. Applying the law of cosines, the first condition yields $\langle p, q \rangle \geq 1 - \frac{r^2}{2}$ and the second condition yields $\langle p, q \rangle \leq 1 - \frac{r^2}{2} - \frac{2\epsilon r^2 + (\epsilon r)^2}{2} < 1 - \frac{r^2}{2} - \frac{\epsilon}{25}$. So, it suffices to take $c = 1/25$.

Now suppose that $r < 0.2$. For each random vector v we have that $\mathbb{E}[\text{sign}(X_i^T v \cdot X_j^T v)] = 1 - \frac{2\theta(X_i, X_j)}{\pi}$, where $\theta(X_i, X_j)$ denotes the angle between X_i, X_j . Since expectations of independent random variables multiply, we have that, for each k ,

$$\mathbb{E}[z_{k,i} z_{k,j}] = (1 - 2 \cdot \theta(X_i, X_j)/\pi)^q.$$

Now let $\theta_r = \cos^{-1}(1 - r^2/2)$,

$$\|X_i - X_j\| \leq r \implies \theta(X_i, X_j) \leq \theta_r \implies \mathbb{E}[\langle Z_i, Z_j \rangle] \geq d(1 - 2\theta_r/\pi)^q$$

$$\|X_i - X_j\| \geq (1 + \epsilon)r \implies \theta(X_i, X_j) \geq (1 + \epsilon/2)\theta_r \implies \mathbb{E}[\langle Z_i, Z_j \rangle] \leq d(1 - 2(1 + \epsilon/2)\theta_r/\pi)^q.$$

Notice that,

$$\frac{(1 - 2(1 + \epsilon/2)\theta_r/\pi)^q}{(1 - 2\theta_r/\pi)^q} < 1 - \epsilon/10,$$

for $q = \lfloor \pi/(2\theta_r) \rfloor$ and since $n^{-0.9} \leq r \leq 0.2$. Notice that $d(1 - 2\theta_r/\pi)^q \in [0.3d, 0.5d]$. Hence, if $\|X_i - X_j\| \leq r$ and $\|X_l - X_k\| \geq (1 + \epsilon)r$,

$$\mathbb{E}[\langle Z_l, Z_k \rangle] < (1 - \epsilon/10)\mathbb{E}[\langle Z_i, Z_j \rangle] \leq \mathbb{E}[\langle Z_i, Z_j \rangle] - 0.3d\epsilon/10,$$

By a union bound over Chernoff bounds, since $d = \log^3 n$, with probability $1 - o(1/\text{poly}(n))$, the inner products between any two columns of Z differs from their expectations by $o(d)$. After performing the scaling procedure, and due to the fact that $d(1 - 2\theta_r/\pi)^q \leq 0.5d$, we conclude that it suffices to compute **Inner Product ApprxNet** with $\rho = (1 - 2\theta_r/\pi)^q$ and approximation error $\epsilon/100$.

The runtime of all components of the algorithm aside from the calls to **Inner Product ApprxNet** is bounded by $\tilde{O}(n/\cos^{-1}(1 - r^2/2)) = \tilde{O}(n^{1.9})$. \square

F Proof of Theorem 14

We present a randomized approximation algorithm which, given a pointset in \mathbb{R}^d and distance parameter r , returns the points that have at least one neighbor at distance at most r .

DelFar

Input: Matrix $X = [x_1, \dots, x_n]$ with each $x_i \in \mathbb{R}^d$, parameter $r \in \mathbb{R}$, constant $\epsilon \in (0, 1/2]$.

Output: $F' \subseteq \{x_1, \dots, x_n\}$.

- Let Y, r' be the output by algorithm **Standardize** on input X, r with parameter $\epsilon/4$.
- If $r \geq 1/n^{0.9}$ run **ApprxNet(Large radius)** on input $Y, \epsilon/4, r$ and return points which correspond to the set $F' \leftarrow X \setminus F$.
- If $r < 1/n^{0.9}$ run **ApprxNet(Small radius)** on input $Y, \epsilon/4, r$ and return points which correspond to the set $F' \leftarrow X \setminus F$.

By Theorems 11, 12, 9, both **ApprxNet(Large radius)** and **ApprxNet(Small radius)** return a set F , the subset of the centers of r -net that are isolated, i.e. the points that do not have any neighbor at distance $(1 + \epsilon)r$. Also, both procedures run in $\tilde{O}(dn^{2-\Theta(\sqrt{\epsilon})})$. Thus, **DelFar** on input a $d \times n$ matrix X , a radius $r \in \mathbb{R}$ and a fixed constant $\epsilon \in (0, 1/2]$ returns a set $F' \subseteq \{x_1, \dots, x_n\}$, which contains all the points (vectors) of X that have at least one neighbor at distance r . Additionally, the algorithm costs $\tilde{O}(dn^{2-\Theta(\sqrt{\epsilon})})$ time and succeeds with probability $1 - o(n^{0.04})$.

G A general framework for high dimensional distance problems

In this section, we modify a framework originally introduced by [HR15], which provides an efficient way for constructing approximation algorithms for a variety of well known distance problems. We present the algorithm **Net and Prune** of [HR15], modified to call the algorithms **ApprxNet** and **DelFar**. We claim that this algorithm computes, with high probability, a constant spread interval and costs $O(dn^{1.999999})$ time.

We assume the existence of a fast approximate decider procedure for the problems we want to address using this framework, specifically an algorithm that runs in $\tilde{O}(dn^{2-\Theta(\sqrt{\epsilon})})$, where ϵ is the approximation factor. Formally,

Definition 23. *Given a function $f : X \rightarrow \mathbb{R}$, we call a decider procedure a $(1 + \epsilon)$ -decider for f , if for any $x \in X$ and $r > 0$, $\text{decider}(r, x)$ returns one of the following: (i) $f(x) \in [\alpha, (1 + \epsilon)\alpha]$, where α is some real number, (ii) $f(x) < r$, or (iii) $f(x) > r$.*

Additionally, we assume the problems we seek to improve with this method have the following property: if the decider returns that the optimal solution is smaller than a fixed value r , we can efficiently remove all points that do not have any neighbor at distance at most r and this does not affect the optimal solution. Let us denote $f(X)$ the optimal solution of a problem for input X .

Net & Prune

Input: An instance (X, Γ) s.t. $X \subseteq \mathbb{R}^d$.

Output: An interval $[x, y]$ containing the optimal value.

- $X_0 = X, i = 0$
- While TRUE do

- Choose at random a point $x \in X_i$ and compute its nearest neighbor distance, l_i
- Call $\frac{3}{2}$ -decider($2l_i/3, X_i$) and $\frac{3}{2}$ -decider(cl_i, X_i). Do one of the following:
 - * If $\frac{3}{2}$ -decider($2l_i/3, X_i$) returns $f(X_i) \in [x, y]$, return $f(X) \in [x/2, 2y]$
 - * If $\frac{3}{2}$ -decider(cl_i, X_i) returns $f(X_i) \in [x', y']$, return $f(X) \in [x'/2, 2y']$
 - * If $2l_i/3$ is too small and cl_i too large, return $[l_i/3, 2cl_i]$
 - * If $2l_i/3$ is too large, call $X_{i+1} = \text{DelFar}(2l_i/3, X_i, \frac{3}{2})$
 - * If cl_i is too small, $X_{i+1} = \text{ApprxNet}(4l_i, X_i, \frac{3}{2})$
- $i = i + 1$,

Let us denote as $|X_i^{\leq l}|$ and $|X_i^{\geq l}|$ the set of points in X , whose nearest neighbor distance is smaller than l and greater than l , respectively.

Theorem 24. *Assume that the DelFar algorithm and the ApprxNet algorithm succeed with probability $1 - \frac{1}{n^{0.01}}$. The algorithm Net & Prune (X, Γ) runs in expected $O(dn^{1.999999})$ time.*

Proof. In each iteration of the while loop the algorithm calls on input X_i the $\frac{3}{2}$ -decider procedure and either ApprxNet or DelFar, all of which cost $O(d|X_i|^{1.999999})$ time. Thus, the total running time of the algorithm is $O(\sum_{i=0}^{i=k-1} d|X_i|^{1.999999})$, where k denotes the last iteration of the while loop.

In the $(i+1)$ th iteration of the while loop, where $(i+1 < k)$, let's assume that x_1, x_2, \dots, x_m is the points' labels in increasing order of their nearest neighbor distance in X_i . If j is the index of the chosen point on the first step of the algorithm and $X_i^{\geq j}$ and $X_i^{\leq j}$ are the subsets of points with index $\geq j$ and $\leq j$, respectively, then we call i a successful iteration when $j \in [m/4, 3m/4]$. Then, we have that $|X_i^{\geq j}| \geq |X_{i+1}|/4$ and $|X_i^{\leq j}| \geq |X_{i+1}|/4$ for a successful iteration. The probability that $i+1$ is a successful iteration is $1/2$.

At each iteration, but the last, either ApprxNet or DelFar gets called. Thus, for any successful iteration, a constant fraction of the point set is removed (it follows from Lemma 3.2.3 in [HR15] and Theorem 14). Also, the algorithms $(1 + \epsilon)$ -decider, ApprxNet and DelFar succeed at every call with probability $1 - \frac{O(\log n)}{n^{0.01}} = 1 - o(1)$, since the expected number of iterations is $O(\log n)$. Hence, the expected running time of the algorithm is $O(dn^{1.999999})$, given the above algorithms succeed. \square

At every step, either far points are being removed or we net the points. If the DelFar algorithm is called, then with small probability we remove a point which is not far. This obviously affects the optimal value, thus we will prove the correctness of the algorithm with high probability. On the other hand, if the ApprxNet algorithm is called, the net radius is always significantly smaller than the optimal value, so the accumulated error in the end, which is proportional to the radius of the last net computation, is also much smaller than the optimal value. For the following proofs we assume both DelFar and ApprxNet algorithms succeed, which occurs with probability $1 - o(1)$.

Lemma 25. *For every iteration i , we have $|f(X_i) - f(X_0)| \leq 16l_i$.*

Proof. Let I be the set of indices of the ApprxNet iterations up to the i th iteration. Similarly, let I' be the set of iterations where DelFar is called.

If ApprxNet was called in the j th iteration, then X_j is at most a $6l_j$ -drift of X_{j-1} , therefore $|f(X_j) - f(X_{j-1})| \leq 12l_j$. Also, if DelFar is called in the j th iteration, then $f(X_j) = f(X_{j-1})$ (by Theorem 14). Let $m = \max I$, we have that,

$$|f(X_i) - f(X_0)| \leq \sum_{j=1}^i |f(X_j) - f(X_{j-1})| = \sum_{j \in I} |f(X_j) - f(X_{j-1})| + \sum_{j \in I'} |f(X_j) - f(X_{j-1})|$$

$$\leq \sum_{j \in I} 12l_j + \sum_{j \in I'} 0 \leq 12l_m \sum_{j=0}^{\infty} \left(\frac{1}{4}\right)^j \leq 16l_m \leq 16l_i$$

,where the second inequality holds since for every $j < i$, in the beginning of the j th iteration of the while loop, the set of points X_{j-1} is a subset of the net points of a $4l_i$ -net, therefore $l_j \geq 4l_i$. \square

Lemma 26. *For any iteration i of the while loop such that **ApprxNet** gets called, we have $l_i \leq f(X_0)/\eta$, where $\eta = c - 16$.*

Proof. We will prove this with induction. Let m_1, m_2, \dots, m_t be the indices of the iterations of the while loop in which **ApprxNet** gets called.

Base: In order for **ApprxNet** to get called we must have $\eta l_{m_1} < cl_{m_1} < f(X_{m_1-1})$ and since this is the first time **ApprxNet** gets called we have $f(X_{m_1-1}) = f(X_0)$. Therefore, $\eta l_{m_1} < f(X_0)$.

Inductive step: Suppose that $l_{m_j} \leq f(X_0)/\eta$, for all $m_j < m_i$. If a call to $\frac{3}{2}$ -rNet is made in iteration m_i then again $cl_{m_i} < f(X_{(m_i)-1}) = f(X_{m_i-1})$. Thus, by the induction hypothesis and Lemma 25 we have,

$$l_{m_i} < \frac{f(X_{m_i-1})}{c} \leq \frac{f(X_0) + 16l_{m_i-1}}{c} \leq \frac{f(X_0) + 16f(X_0)/\eta}{c} = \frac{1 + 16/\eta}{c} f(X_0) = f(X_0)/\eta$$

\square

Therefore, if we set $c = 64$ we have $\eta = 48$, thus by Lemma 25 and Lemma 26,

$$|f(X_i) - f(X_0)| \leq 16l_i \leq 16f(X_0)/\eta = f(X_0)/3$$

Corollary 27. *For $c \geq 64$ and for any iteration i we have:*

- $(2/3)f(X_0) \leq f(X_i) \leq (4/3)f(X_0)$,
- if $f(X_i) \in [x, y]$, then $f(X_0) \in [(3/4)x, (3/2)y] \subseteq [x/2, 2y]$,
- if $f(X_0) > 0$ then $f(X_i) > 0$.

Theorem 28. *For $c \geq 64$, the Net & Prune algorithm computes in $O(dn^{1.999999})$ time a constant spread interval containing the optimal value $f(X)$, with probability $1 - o(1)$.*

Proof. Consider the iteration of the while loop at which Net & Prune terminates. If the interval $[x, y]$ was computed by the $\frac{3}{2}$ -decider, then it has spread $\leq \frac{3}{2}$. Thus, by Corollary 27 the returned interval $[x', y'] = [x/2, 2y]$ contains the optimal value and its spread is ≤ 6 . Similarly, if $2l_i/3$ is too small and cl_i too large, then the returned interval is $[\frac{l_i}{3}, 2cl_i]$ and its spread is 384. \square

G.1 Proof of Theorem 16

Proof. For this particular problem, the optimal solution is not affected by the **DelFar**'s removal of the points with no other point at distance at most r . Also, each time the **ApprxNet** algorithm is called, for a fixed distance r , the drift of the optimal solution is at most $2r$. Thus, Theorem 28 holds, and we compute a constant spread interval $[x, y]$ containing the optimal value, with high probability. We then apply binary search on values $x, (1 + \epsilon)x, (1 + \epsilon)^2x, \dots, y$ using the algorithm k th NND Decider. We perform $O(1/\log(1 + \epsilon)) = O(1/\epsilon^2)$ iterations, hence the total amount of time needed is $\tilde{O}(dn^{2-\Theta(\sqrt{\epsilon})})$ and the algorithm succeeds with high probability $1 - o(1)$. \square