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Approximation and Parameterized Algorithms for Segment Set Cover

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Master's thesis

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9

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10 **Supervisor's statement**

11 Hereby I confirm that the presented thesis was prepared under my supervision and
12 that it fulfils the requirements for the degree of Master of Computer Science.

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Abstract

23 In this thesis we analyze approximation and parametrization of GEOMETRIC SET COVER
 24 with segments. To achieve interesting results we define a concept of GEOMETRIC SET COVER
 25 with δ -extension, where we need to cover the points by geometric objects, which are extended
 26 by a tiny fraction, but we compare the solution size to the optimum solution without ex-
 27 tension. We prove that SEGMENT SET COVER is APX-hard even if we restrict segments to
 28 be axis-parallel and allow $\frac{1}{2}$ -extension (Chapter 3). We provide an FPT algorithm for un-
 29 weighted SEGMENT SET COVER parametrized by the size of the solution k (Section 4.1) and
 30 for WEIGHTED SEGMENT SET COVER with δ -extension (Section 4.2). Finally, we prove that
 31 WEIGHTED SEGMENT SET COVER is W[1]-hard and there does not exist an algorithm run-
 32 ning in time $f(k) \cdot n^{o(\sqrt{k})}$ solving this problem even if we restrict the segments to 3 directions
 33 (Chapter 5).

34

Keywords

35 set cover, geometric set cover, weighted set cover, FPT, W[1]-hard, APX-hard, ETH, grid
 36 tiling, MAX-(3,3)-SAT, δ -extension

37

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43 Theory of computation \rightarrow Design and analysis of algorithms \rightarrow Approximation algorithms
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45

46

Tytuł pracy w języku polskim

47 Algoritmy aproksymacyjne i parametryzowane dla problemu pokrywania punktów
 48 odcinkami na płaszczyźnie

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Chapter 1

Introduction

Some problems in Computer Science are known to be NP-complete, meaning that assuming $P \neq NP$ there is no polynomial time algorithm that can solve these problems. Even so, they still can be amenable to different approaches, such as approximation or parameterization.

Definition 1.1. In the **SET COVER** problem we are given a set of elements (universe) \mathcal{C} and a family of sets \mathcal{P} that are subsets of the universe \mathcal{C} and sum up to the whole \mathcal{C} . Our task is to find a set $\mathcal{R} \subseteq \mathcal{P}$ such that $\bigcup \mathcal{R} = \mathcal{C}$ and the size of \mathcal{R} is minimum possible.

SET COVER is a classical example of an NP-complete problem, which has been proven in [Dinur and Steurer, 2014] to be inapproximable with factor $(1 - o(1)) \ln n$ assuming $P \neq NP$ (which is a stronger result than APX-hardness), and W[2]-complete with the natural parameterization, see Theorem 13.21 in [Cygan et al., 2015]. However restricting the problem to various specialized settings can lead to more tractable special cases. In this thesis we take a closer look at the GEOMETRIC SET COVER problem in the plane, where elements to cover are points in the plane and sets to cover them with are geometric objects.

Definition 1.2. **SEGMENT SET COVER** is GEOMETRIC SET COVER where objects that we cover the points with are segments in the plane.

Approximation Over the years there has been a lot of work related to approximation algorithms for GEOMETRIC SET COVER. Notably, GEOMETRIC SET COVER with unweighted unit disks admits a PTAS (see Corollary 1.1 in [Mustafa and Ray, 2010]). When we consider the same problem with weighted unit disks (or unit squares), the problem admits a QPTAS [Mustafa et al., 2014], see also [Pilipczuk et al., 2020]. On the other hand, [Chan and Grant, 2014] proves that GEOMETRIC SET COVER with unweighted axis-parallel fat rectangles is APX-hard; they also show similar hardness for GEOMETRIC SET COVER with many other standard geometric objects.

Parameterization We consider GEOMETRIC SET COVER parameterized by the size of solution. GEOMETRIC SET COVER with unit squares was first proven to be W[1]-hard in [Marx, 2005] (Theorem 5). A later follow-up work [Marx and Pilipczuk, 2015] shows that there is an algorithm running in time $n^{\mathcal{O}(\sqrt{k})}$ that solves GEOMETRIC SET COVER with unit squares or disks and that there is no algorithm running in time $f(k) \cdot n^{o(\sqrt{k})}$ for any computable f under the Exponential-Time Hypothesis, so this is a tight bound for this problem.

We also consider parameterization of weighted problems. There does not seem to be a consensus of what parameterization in the weighted setting is exactly; there was an at-

tempt to introduce a quite complicated general framework of weighted parameterized setting in [Shachnai and Zehavi, 2017]. Kernels for several well-known weighted problems such as WEIGHTED SUBSET SUM or WEIGHTED KNAPSACK are presented in [Etscheid et al., 2017]. Another work [Kim et al., 2021] considers weighted parameterization of WEIGHTED DIRECTED FEEDBACK SET and WEIGHTED *st*-CUT.

δ -extension In this paper, we focus on SEGMENT SET COVER with δ -extension. δ -extension is a problem relaxation method based on the δ -shrinking model which was introduced in [Adamaszek et al., 2015] to provide an interesting result for the MAXIMUM WEIGHT INDEPENDENT SET OF RECTANGLES problem. In this problem one needs to find a set of non-overlapping weighted rectangles with maximum sum of weight possible. In the δ -shrinking relaxed problem the returned set of rectangles must be non-overlapping after all the rectangles are shrunk by a tiny fraction δ towards the centre of symmetry. This problem is easier, because we compare this result to the optimum result before the shrinking. It might even lead to finding a set with result better than the optimum for the original problem. The author in [Adamaszek et al., 2015] presents a PTAS for MAXIMUM WEIGHT INDEPENDENT SET OF RECTANGLES with δ -shrinking, which is later improved to EPTAS in [Pilipczuk et al., 2016] alongside presenting a new FPT result for this problem with the natural parameterization. Later the similar δ -shrinking model was used in [Wiese, 2018] to present a PTAS for MAXIMUM WEIGHT INDEPENDENT SET OF POLYGONS with δ -shrinking.

Definition 1.3. For any $\delta > 0$ and a centre-symmetric convex object L with centre of symmetry $S = (x_s, y_s)$, the **δ -extension** of L is the open set of points:

$$L^{+\delta} = \text{Int}\{(1 + \delta) \cdot (x - x_s, y - y_s) + (x_s, y_s) : (x, y) \in L\},$$

where Int denotes the interior of a set of points. That is, $L^{+\delta}$ is interior of the image of L under homothety centred at S with scale $(1 + \delta)$.

Analogous to δ -shrinking, δ -extension provides a framework for relaxing GEOMETRIC SET COVER problems, where we allow the returned set of objects \mathcal{R} to *almost* cover the points in the universe by requiring that they are covered by \mathcal{R} after δ -extension, i.e. by set $\mathcal{R}^{+\delta}$. The same concept could be used for GEOMETRIC SET HITTING problems.

For a longer discussion of this concept see Section 2.4.

Similar model is used to prove that GEOMETRIC SET COVER with fat polygons relaxed with δ -extension admits EPTAS [Har-Peled and Lee, 2009]. The δ -extension model presented there is well-defined only for fat polygons. It extends the object by all the points that have distance to the closest point in the object P no larger than $\delta \cdot \text{rad}(P)$, where $\text{rad}(P)$ is a radius of a circle inscribed into P . Since segments do not have any circle inscribed into them, the definition presented there cannot be utilized for this problem. Polygons extended by δ -extension defined in Definition 1.3 covers a superset of set of points that object extended by δ -extension defined in [Har-Peled and Lee, 2009]. Since our definition is more permissive for any polygon, the EPTAS from [Har-Peled and Lee, 2009] also works for polygons extended by our δ -extension.

Our contribution

In this paper we make the following contributions.

We show that approximation of SEGMENT SET COVER, even if segments are axis-parallel and we relax the problem with $\frac{1}{2}$ -extension, is APX-hard (Theorem 1.1).

151 **Theorem 1.1.** (**SEGMENT SET COVER with axis-parallel segments and $\frac{1}{2}$ -extension**
 152 **is APX-hard**). SEGMENT SET COVER with axis-parallel segments in the 2D plane is APX-
 153 hard (even with $\frac{1}{2}$ -extension). That is, assuming $P \neq NP$, there does not exist a PTAS for
 154 this problem.

155 Theorem 1.1 implies the following. Note that segments are just degenerated rectangles.

156 **Corollary 1.1.** (**GEOMETRIC SET COVER with rectangles is APX-hard**). GEOMET-
 157 RIC SET COVER with axis-parallel rectangles is APX-hard (even with $\frac{1}{2}$ -extension).

158 This expands the previous result of [Chan and Grant, 2014] that GEOMETRIC SET COVER
 159 with axis-parallel fat rectangles is APX-hard, we improved the result that rectangles no longer
 160 have to be fat (Corollary 1.1). This also proves that the assumption in [Har-Peled and Lee,
 161 2009] for EPTAS about polygons being fat is necessary, because cover with arbitrary polygons
 162 with δ -extension is APX-hard.

163 We also provide two FPT algorithms for parameterized SEGMENT SET COVER (Theo-
 164 rem 1.2) and WEIGHTED SEGMENT SET COVER relaxed with δ -extension (Theorem 1.3).

165 **Theorem 1.2.** (**FPT for SEGMENT SET COVER**). There exists an algorithm that given
 166 a family \mathcal{P} of segments (in any direction), a set of points \mathcal{C} and a parameter k , runs in time
 167 $k^{O(k)}(|\mathcal{C}| \cdot |\mathcal{P}|)^2$, and outputs a solution $\mathcal{R} \subseteq \mathcal{P}$ such that $|\mathcal{R}| \leq k$ and \mathcal{R} covers all points
 168 in \mathcal{C} , or determines that such a set \mathcal{R} does not exist.

169 **Theorem 1.3.** (**FPT for WEIGHTED SEGMENT SET COVER with δ -extension**). There
 170 exists an algorithm that given a family \mathcal{P} of n weighted segments (in any direction), a set of
 171 m points \mathcal{C} , and parameters k and $\delta > 0$, runs in time $f(k, \delta) \cdot (nm)^c$ for some computable
 172 function f and a constant c and outputs a set \mathcal{R} such that:

- 173 • $\mathcal{R} \subseteq \mathcal{P}$,
- 174 • $|\mathcal{R}| \leq k$,
- 175 • $\mathcal{R}^{+\delta}$ covers all points in \mathcal{C} ,
- 176 • the weight of \mathcal{R} is not greater than the weight of an optimum solution of size at most k
 177 for this problem without δ -extension,

178 or determines that there is no set \mathcal{R} with $|\mathcal{R}| \leq k$ such that \mathcal{R} covers all points in \mathcal{C} .

179 On the other hand, we prove that WEIGHTED SEGMENT SET COVER is W[1]-hard even
 180 if segments are limited to 3 directions (Theorem 1.4) and assuming ETH there does not
 181 exist algorithm for this problem that runs in time $f(k)(|\mathcal{C}| + |\mathcal{P}|)^{O(\sqrt{k})}$. See Figure 1.1 for a
 182 summary of parameterized results for WEIGHTED SEGMENT SET COVER. Similar table for
 183 unweighted problem is present in Figure 1.2.

184 **Theorem 1.4.** Consider the problem of covering a set \mathcal{C} of points by selecting at most k
 185 segments from a set of segments \mathcal{P} with non-negative weights $w : \mathcal{P} \rightarrow \mathbb{R}^+$ so that the weight
 186 of the cover is minimal. Then this problem is W[1]-hard when parameterized by k and assuming
 187 ETH, there is no algorithm for this problem with running time $f(k) \cdot (|\mathcal{C}| + |\mathcal{P}|)^{O(\sqrt{k})}$ for any
 188 computable function f . Moreover, this holds even if all segments in \mathcal{P} are axis-parallel or
 189 right-diagonal.

See Section 2.1 for exact definitions of axis-parallel and right-diagonal segments.

This result is particularly interesting, because problem without weights is FPT and weighted problem is W[1]-hard. Moreover δ -extension allowed us to provide an FPT algorithm for a problem, which is W[1]-hard otherwise.

Note that the result of Theorem 1.4 is not tight: there exists a simple algorithm running in time $\mathcal{O}(f(k)(|\mathcal{C}| + |\mathcal{P}|)^k)$. So the question whether there exists an algorithm for this problem running in time $f(k) \cdot (|\mathcal{C}| + |\mathcal{P}|)^{o(k)}$ is still open.

Permissive FPT is a relaxed FPT problem, where we need to find solution of *any* size in FPT-time, but we compare it to the optimum solution of size at most k . Idea for permissive FPT in local search was presented in [Marx and Schlotter, 2011], [Gaspers et al., 2012].

Theorem 1.4 can be improved to show that a permissive FPT algorithm does not exist. This is formulated precisely in Theorem 5.2.

	exact	δ -extension
axis-parallel	?	FPT*
3 directions	W[1]-hard	FPT*
any direction	W[1]-hard*	FPT

Figure 1.1: Our results for WEIGHTED SEGMENT SET COVER parameterized by the size of solution. Results marked with * are not explicitly given in this thesis, but they trivially follow from stronger results shown in the other cells of the table.

	exact	δ -extension
axis-parallel	FPT*	FPT*
3 directions	FPT*	FPT*
any direction	FPT	FPT*

Figure 1.2: Our results for unweighted SEGMENT SET COVER parameterized by the size of solution. Results marked with * are not explicitly given in this thesis, but they trivially follow from stronger results shown in the other cells of the table.

Future work. There are two aforementioned problems that relate to Theorem 1.4 and were not solved in this thesis. We have not presented W[1]-hardness proof of WEIGHTED SEGMENT SET COVER where segments are limited to 3 directions, but it may be possible to improve this construction to use segments in 2 directions instead of 3 directions. The other question is what is the tight bound for this problem. The simple algorithm solving this problem is running in time $\mathcal{O}(f(k)(|\mathcal{C}| + |\mathcal{P}|)^k)$.

Another problem to consider is whether GEOMETRIC SET HITTING relaxed with δ -extension can yield some better results.

210 Chapter 2

211 Preliminaries

212 In this chapter we present some basic definitions that will be used later.

213 2.1. Geometric Set Cover

214 Whenever speaking about GEOMETRIC SET COVER, we consider it in the 2-dimensional
215 plane.

216 In the GEOMETRIC SET COVER problem we are given \mathcal{P} — a set of objects, which
217 are connected subsets of the plane and \mathcal{C} — a set of points in the plane. The task is to choose
218 $\mathcal{R} \subseteq \mathcal{P}$ such that every point in \mathcal{C} is inside some object from \mathcal{R} and $|\mathcal{R}|$ is minimized. We
219 will mostly consider the case where \mathcal{P} consists of segments in the plane.

220 In the weighted setting, there is some given weight function $f : \mathcal{P} \rightarrow \mathbb{R}^+$ and we would
221 like to find a solution \mathcal{R} that minimizes $\sum_{R \in \mathcal{R}} f(R)$.

222 **Definition 2.1.** A segment is **axis-parallel** if it lies on a line that is either horizontal $y = c$
223 or vertical $x = c$.

224 **Definition 2.2.** A line is **right-diagonal** if it is described by the linear function $x + y = d$
225 for some $d \in \mathbb{R}$. Segment is **right-diagonal** if its direction is a right-diagonal line.

226 2.2. Parameterization

227 In the parameterized setting of the GEOMETRIC SET COVER for a given k , our task is to
228 either find a solution \mathcal{R} such that $|\mathcal{R}| \leq k$ or decide that there is no such solution.

229 **Definition 2.3.** A **Fixed-parameter Tractable (FPT)** algorithm for a problem with pa-
230 rameter k and instance size n is an algorithm running in time $f(k) \cdot n^c$ for some constant c
231 and some computable function f .

232 **Definition 2.4.** Boolean formula is in **conjunctive normal form (CNF)** if it is a con-
233 junction of one or more formulas, which are disjunction of literals. **k -CNF** formula is a CNF
234 formula, where every disjunction consists of at most k literals.

235 **Definition 2.5.** **k -SAT** problem is a boolean satisfiability problem of k -CNF formulas.
236 Given k -CNF formula, one must answer if there exists any variable assignment that satisfies
237 the formula.

Definition 2.6. For $k \geq 3$ set us define S_k as a set of constants σ such that there exists an algorithm solving k -SAT running in time $\mathcal{O}^*(2^{\sigma n})$. Let s_k be the infimum of the set S_k .

Exponential Time Hypothesis (ETH) asserts that $s_3 > 0$. This conjecture implies that there does not exist an algorithm solving 3-SAT running in time $2^{o(n)}$.

To see the definition of a $W[1]$ -hard problem and W hierarchy, see Chapter 13.3 of [Cygan et al., 2015]. When proving that a problem is $W[1]$ -hard, we are going to use Theorem 5.1, which was proved in [Marx, 2007].

2.3. Approximation

Let us recall some definitions related to optimization problems.

Definition 2.7. A **polynomial-time approximation scheme (PTAS)** for a minimization problem Π is a family of algorithms \mathcal{A}_ϵ for every $\epsilon > 0$ such that \mathcal{A}_ϵ takes an instance I of Π and in polynomial time finds a solution that is within a factor of $(1 + \epsilon)$ of being optimal. This means that the reported solution has weight at most $(1 + \epsilon)\text{opt}(I)$, where $\text{opt}(I)$ is the weight of an optimal solution to I .

Definition 2.8. A problem Π is **APX-hard** if assuming $P \neq NP$, there exists $\epsilon > 0$ such that there is no polynomial-time $(1 + \epsilon)$ -approximation algorithm for Π .

2.4. δ -extension

Another idea presented here, which can be utilized only when considering the problems with geometric objects, is δ -extension. We define it specifically for the GEOMETRIC SET COVER problem with convex centre-symmetric objects.

Intuitively, we consider a problem with slightly larger objects, which makes the instance more permissive. However, we aim to find a solution that is not larger than the optimum solution to the original problem, so this is substantially easier than just solving the problem for the larger objects. It may even be the case that we are able to find a solution of size smaller than the optimum solution to the original problem.

Formal definition of δ -extended objects is present in Definition 1.3.

The GEOMETRIC SET COVER with δ -extension is a version of GEOMETRIC SET COVER with the following modifications.

- We need to cover all the points in \mathcal{C} by selecting objects from $\{P^{+\delta} : P \in \mathcal{P}\}$ (which always include no fewer points than the objects before δ -extension).
- We look for a solution that is not larger than the optimum solution to the original problem. Note that it does not need to be an optimal solution in the modified problem.

Formally, we have the following.

Definition 2.9. The **GEOMETRIC SET COVER problem with δ -extension** is the problem where for an input instance $I = (\mathcal{P}, \mathcal{C})$ of GEOMETRIC SET COVER, the task is to output a solution $\mathcal{R} \subseteq \mathcal{P}$ such that the δ -extended set $\{R^{+\delta} : R \in \mathcal{R}\}$ covers \mathcal{C} and is not larger than the optimal solution to the problem without extension, i.e. $|\mathcal{R}| \leq |\text{opt}(I)|$.

At last, we formulate a definition of the polynomial-time approximation scheme (PTAS) for a problem with δ -extension.

277 **Definition 2.10.** A PTAS for GEOMETRIC SET COVER with δ -extension is a family of
 278 algorithms $\{\mathcal{A}_{\delta,\epsilon}\}_{\delta,\epsilon>0}$ that each takes as an input instance $I = (\mathcal{P}, \mathcal{C})$ of GEOMETRIC SET COVER
 279 where objects are centre-symmetric and convex, and in polynomial-time outputs a solution
 280 $\mathcal{R} \subseteq \mathcal{P}$ such that the δ -extended set $\{R^{+\delta} : R \in \mathcal{R}\}$ covers \mathcal{C} and is within a $(1 + \epsilon)$ factor of
 281 the optimal solution to this problem without extension, i.e. $(1 + \epsilon)|\mathcal{R}| \leq |\text{opt}(I)|$.

282 2.5. Weighted Geometric Set Cover

283 In this thesis we also consider a WEIGHTED GEOMETRIC SET COVER problem, which is a
 284 combination of the weighted and parameterized setting described in Section 2.1. We already
 285 argued in the introduction that there is no consensus of how it is defined, but when we discuss
 286 the weighted parameterized setting we will consider the following definition. There is a given
 287 weight function $f : \mathcal{P} \rightarrow \mathbb{R}^+$ and we would like to find a solution \mathcal{R} such that $|\mathcal{R}| \leq k$ and
 288 $\sum_{R \in \mathcal{R}} f(R)$ is minimum possible among such sets \mathcal{R} .

289 **Definition 2.11.** The WEIGHTED GEOMETRIC SET COVER problem with δ -extension
 290 is the problem where for an input instance $I = (\mathcal{P}, \mathcal{C}, f)$ of WEIGHTED GEOMETRIC SET COVER,
 291 the task is to output a solution $\mathcal{R} \subseteq \mathcal{P}$ such that the δ -extended set $\{R^{+\delta} : R \in \mathcal{R}\}$ covers
 292 \mathcal{C} and it has weight not larger than the optimal solution to the problem without extension,
 293 i.e. $\sum_{R \in \mathcal{R}} f(R) \leq |\text{opt}(I)|$.

294 We also consider weighted parameterized setting with δ -extension, which we formally
 295 define below.

296 **Definition 2.12.** The WEIGHTED GEOMETRIC SET COVER problem with δ -extension
 297 parameterized by the size of a solution is a problem where for an input instance
 298 $I = (\mathcal{P}, \mathcal{C}, f, k)$ of WEIGHTED GEOMETRIC SET COVER parameterized by the size of a so-
 299 lution k , the task is to output a solution $\mathcal{R} \subseteq \mathcal{P}$ such that the δ -extended set $\{R^{+\delta} : R \in \mathcal{R}\}$
 300 covers \mathcal{C} , uses no more than k sets, i.e. $|\mathcal{R}| \leq k$ and it has weight not larger than the optimal
 301 solution to the problem without extension, i.e. $\sum_{R \in \mathcal{R}} f(R) \leq |\text{opt}(I)|$.

Chapter 3

APX-hardness of geometric set cover

In this section we analyze whether there exists a PTAS for geometric set cover for rectangles. We show that Segment Set Cover is APX-hard even if we can restrict this problem to a very simple setting: segments parallel to axes and allow $\frac{1}{2}$ -extension.

Our result can be summarized in the following theorem and this section aims to prove it.

Theorem 1.1. (**SEGMENT SET COVER with axis-parallel segments and $\frac{1}{2}$ -extension is APX-hard**). SEGMENT SET COVER with axis-parallel segments in the 2D plane is APX-hard (even with $\frac{1}{2}$ -extension). That is, assuming $P \neq NP$, there does not exist a PTAS for this problem.

We prove Theorem 1.1 by taking a problem that is APX-hard and showing a reduction. For this problem we choose MAX-(3,3)-SAT which we define below.

3.1. MAX-(3,3)-SAT

See Definition 2.4 for the definition of the k -CNF formula.

Definition 3.1. MAX-3SAT is the following maximization problem. We are given a 3-CNF formula, and we need to find a boolean assignment of variables that satisfies the most clauses.

Definition 3.2. MAX-(3,3)-SAT is a variant of MAX-3SAT with an additional restriction that every variable appears in exactly 3 clauses and every clause contains exactly 3 literals of 3 different variables. Note that thus, the number of clauses is equal to the number of variables.

In our proof of Theorem 1.1 we use hardness of approximation of MAX-(3,3)-SAT proved in [Håstad, 2001] and described in Theorem 3.1 below.

Definition 3.3. MAX-3SAT formula with m clauses is **at most α -satisfiable**, if every assignment of variables satisfies no more than αm clauses.

Theorem 3.1 ([Håstad, 2001]). For any $\epsilon > 0$, it is NP-hard to distinguish satisfiable (3,3)-SAT formulas from at most $(\frac{7}{8} + \epsilon)$ -satisfiable (3,3)-SAT formulas.

3.2. Reduction

The following lemma encapsulates the properties of the reduction described in this section, and it allows us to prove Theorem 1.1.

Lemma 3.1. *Given an instance S of MAX-(3,3)-SAT with n variables and optimum value $\text{opt}(S)$, we can construct an instance $(\mathcal{C}, \mathcal{P})$ of geometric set cover with axis-parallel segments in 2D such that:*

(1) *For every solution to instance S that satisfies k clauses, there exists a solution to $(\mathcal{C}, \mathcal{P})$ of size $15n - k$.*

(2) *For every solution \mathcal{R} to instance $(\mathcal{C}, \mathcal{P})$, there exists a solution to S that satisfies at least $15n - |\mathcal{R}|$ clauses.*

(3) *For every $\mathcal{R} \subseteq \mathcal{P}$, if $\mathcal{R}^{+\frac{1}{2}}$ is a solution to $(\mathcal{C}, \mathcal{P})$, then \mathcal{R} is also a solution to $(\mathcal{C}, \mathcal{P})$.*

Therefore, the optimum size of a solution to $(\mathcal{C}, \mathcal{P})$ is $\text{opt}((\mathcal{C}, \mathcal{P})) = 15n - \text{opt}(S)$.

We prove Lemma 3.1 in subsequent sections. Section 3.3 describes the proposed instance $(\mathcal{C}, \mathcal{P})$. Property (1) is proved by Lemma 3.11, (2) by Lemma 3.13, and finally (3) trivially follows from Lemma 3.10. Firstly let us prove Theorem 1.1 using Lemma 3.1 and Theorem 3.1.

Proof of Theorem 1.1. Consider any $0 < \epsilon < \frac{1}{15.8}$.

Let us assume that there exists a polynomial-time $(1 + \epsilon)$ -approximation algorithm for unweighted geometric set cover with axis-parallel segments in 2D with $\frac{1}{2}$ -extension. We construct an algorithm that solves the problem stated in Theorem 3.1, thereby proving that $P = NP$.

Take an instance S of MAX-(3,3)-SAT to be distinguished and construct an instance of geometric set cover $(\mathcal{C}, \mathcal{P})$ using Lemma 3.1. We now use the $(1 + \epsilon)$ -approximation algorithm for geometric set cover relaxed with $\frac{1}{2}$ -extensions on $(\mathcal{C}, \mathcal{P})$. Denote the size of the solution returned by this algorithm as $\text{approx}^*((\mathcal{C}, \mathcal{P}))$. We prove that if in S one can satisfy at most $(\frac{7}{8} + \epsilon)n$ clauses, then $\text{approx}^*((\mathcal{C}, \mathcal{P})) \geq 15n - (\frac{7}{8} + \epsilon)n$, and if S is satisfiable, then $\text{approx}^*((\mathcal{C}, \mathcal{P})) < 15n - (\frac{7}{8} + \epsilon)n$.

Assume S satisfiable. From the definition of S being satisfiable, we have:

$$\text{opt}(S) = n.$$

From Lemma 3.1 we have:

$$\text{opt}((\mathcal{C}, \mathcal{P})) = 14n.$$

Therefore,

$$\begin{aligned} \text{approx}^*((\mathcal{C}, \mathcal{P})) &\leq (1 + \epsilon)\text{opt}((\mathcal{C}, \mathcal{P})) = 14n(1 + \epsilon) = 14n + 14\epsilon \cdot n = \\ &= 14n + (15\epsilon - \epsilon)n < 14n + \left(\frac{1}{8} - \epsilon\right)n = 15n - \left(\frac{7}{8} + \epsilon\right)n. \end{aligned}$$

Assume S is at most $(\frac{7}{8} + \epsilon)$ satisfiable. From the definition of S being at most $(\frac{7}{8} + \epsilon)n$ satisfiable, we have:

$$\text{opt}(S) \leq \left(\frac{7}{8} + \epsilon\right)n$$

From Lemma 3.1 we have:

$$\text{opt}((\mathcal{C}, \mathcal{P})) \geq 15n - \left(\frac{7}{8} + \epsilon\right)n$$

355 Since a solution to $(\mathcal{C}, \mathcal{P})$ with $\frac{1}{2}$ -extension is also a solution without any extension, by
 356 Lemma 3.1 (3), we have:

$$\text{approx}^*((\mathcal{C}, \mathcal{P})) \geq \text{opt}((\mathcal{C}, \mathcal{P})) = 15n - \left(\frac{7}{8} + \epsilon\right)n$$

357 Therefore, by using the assumed $(1 + \epsilon)$ -approximation algorithm, it is possible to distin-
 358 guish the case when S is satisfiable from the case when it is at most $(\frac{7}{8} + \epsilon)n$ satisfiable: it
 359 suffices to compare $\text{approx}^*((\mathcal{C}, \mathcal{P}))$ with $15n - (\frac{7}{8} + \epsilon)n$. Hence, the assumed approximation
 360 algorithm cannot exist, unless $P = NP$. \square

361 3.3. Construction of the Geometric Set Cover instance

362 We proceed to the proof of Lemma 3.1. That is, we show a reduction from the MAX-(3,3)-
 363 SAT problem to geometric set cover with segments parallel to axes. Moreover, the obtained
 364 instance of geometric set cover will be robust to $\frac{1}{2}$ -extension (have the same optimal solution
 365 after $\frac{1}{2}$ -extension).

366 The construction will be composed of 2 types of gadgets: **VARIABLE-gadgets** and
 367 **CLAUSE-gadgets**. **CLAUSE-gadgets** will be constructed using two **OR-gadgets** connected
 368 together.

369 3.3.1. VARIABLE-gadget

370 **VARIABLE-gadget** is responsible for choosing the value of a variable in a CNF formula. It
 371 allows two minimum solutions of size 3 each. These two choices correspond to the two Boolean
 372 values of the variable corresponding to this gadget.

373 **Points.** Define points a, b, c, d, e, f, g, h as follows, where $L = 22n$:



Figure 3.1: **VARIABLE-gadget**. We denote the set of points marked with black circles as pointsVariable_i , and they need to be covered (are part of the set \mathcal{C}). Note that some of the points are not marked as black dots and exists only to name segments for further reference. We denote the set of red segments as $\text{chooseVariable}_i^{\text{false}}$ and the set of blue segments as $\text{chooseVariable}_i^{\text{true}}$.

374

$$\begin{array}{llll} a := (-3L, 0) & b := (-2L, 0) & c := (-L, 0) & d := (-3L, 1) \\ e := (-2L, 1) & f := (-2L, 2) & g := (L, 0) & h := (L, 2) \end{array}$$

Let us define:

$$\text{pointsVariable} := \{a, b, c, d, e, f\}$$

and, for any $1 \leq i \leq n$,

$$\text{pointsVariable}_i := \text{pointsVariable} + (0, 4i).$$

375 We denote $a_i := a + (0, 4i)$ etc.

376 **Segments.** Let us define:

$$\text{chooseVariable}_i^{\text{true}} := \{(a_i, d_i), (b_i, f_i), (c_i, g_i)\},$$

$$\text{chooseVariable}_i^{\text{false}} := \{(a_i, c_i), (d_i, e_i), (f_i, h_i)\},$$

$$\text{segmentsVariable}_i := \text{chooseVariable}_i^{\text{true}} \cup \text{chooseVariable}_i^{\text{false}}.$$

377 We also name two of these segments for future reference: $\text{xTrueSegment}_i := (c_i, g_i)$,
378 $\text{xFalseSegment}_i := (f_i, h_i)$.

379 **Lemma 3.2.** *For any $1 \leq i \leq n$, points in pointsVariable_i can be covered using 3 segments*
380 *from $\text{segmentsVariable}_i$.*

381 *Proof.* We can use either set $\text{chooseVariable}_i^{\text{true}}$ or $\text{chooseVariable}_i^{\text{false}}$. □

382 **Lemma 3.3.** *For any $1 \leq i \leq n$, points in pointsVariable_i can not be covered with fewer than*
383 *3 segments from $\text{segmentsVariable}_i$.*

384 *Proof.* No segment of $\text{segmentsVariable}_i$ covers more than one point from $\{d_i, f_i, c_i\}$, therefore
385 pointsVariable_i can not be covered with fewer than 3 segments. □

386 **Lemma 3.4.** *For every set $A \subseteq \text{segmentsVariable}_i$ such that A covers pointsVariable_i and*
387 *$\text{xTrueSegment}_i, \text{xFalseSegment}_i \in A$, it holds that $|A| \geq 4$.*

388 *Proof.* No segment from $\text{segmentsVariable}_i$ covers more than one point from $\{a_i, e_i\}$, therefore
389 $\text{pointsVariable}_i - \{c_i, f_i\}$ can not be covered with fewer than 2 segments. □

390 3.3.2. OR-gadget

391 An OR-gadget connects input and output segments (see Figure 3.2) in a way that is supposed
392 to simulate the binary disjunction.

393 Input segments are the only segments that cover points outside of the gadget, as their left
394 ends lie outside of it. Point $v_{i,j}$ is the only one that can be covered by segments that do not
395 belong to the gadget.

396 The OR-gadget has the property that every set of segments that covers all the points in
397 the gadget uses at least 3 segments from it. Moreover, the output segment belongs to the
398 solution of size 3 only if at least one of the input segments belongs to the solution. Therefore,
399 optimum solutions restricted to the OR-gadget behave like a binary disjunction for the input
400 segments.

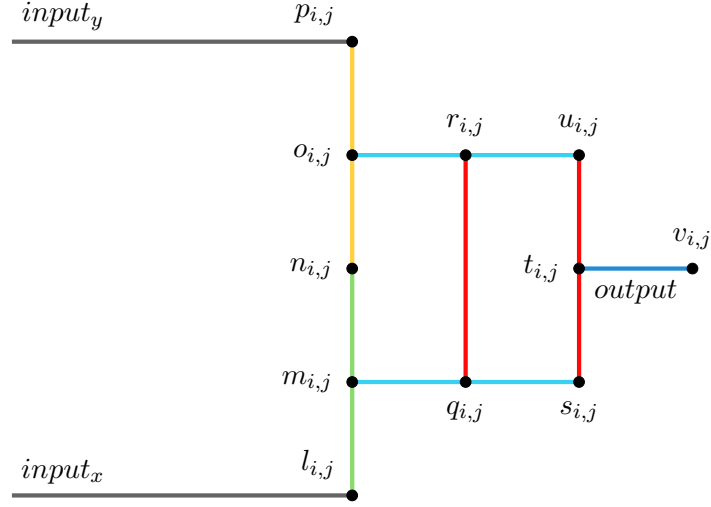


Figure 3.2: **OR-gadget**. Segments from $\text{chooseOr}_{i,j}^{\text{false}}$ are **red**, segments from $\text{chooseOr}_{i,j}^{\text{true}}$ are blue (both **light blue** and **dark blue**), segments from $\text{orMoveVariable}_{i,j}$ are **green** and **yellow**. **Dark blue** segment is the *output* segment. Grey segments input_x and input_y are input segments that are not part of $\text{segmentsOr}_{i,j}$.

401 **Points.** We define

$$\begin{aligned}
 l_0 &:= (0, 0) & m_0 &:= (0, 1) & n_0 &:= (0, 2) & o_0 &:= (0, 3) \\
 p_0 &:= (0, 4) & q_0 &:= (1, 1) & r_0 &:= (1, 3) & s_0 &:= (2, 1) \\
 t_0 &:= (2, 2) & u_0 &:= (2, 3) & v_0 &:= (3, 2)
 \end{aligned}$$

$$vec_{i,j} := (20i + 3 + 3j, 4(n + 1) + 2j)$$

403 For integers i, j , define $\{l_{i,j}, m_{i,j}, \dots, v_{i,j}\}$ as $\{l_0, m_0, \dots, v_0\}$ shifted by $vec_{i,j}$, i.e. $l_{i,j} = l_0 + vec_{i,j}$
 404 etc.

405 Note that $v_{i,0} = l_{i,1}$ (see Figure 3.3). Next, let

$$\text{pointsOr}_{i,j} := \{l_{i,j}, m_{i,j}, n_{i,j}, o_{i,j}, p_{i,j}, q_{i,j}, r_{i,j}, s_{i,j}, t_{i,j}, u_{i,j}\}$$

406 Note that $\text{pointsOr}_{i,j}$ does not include the point $v_{i,j}$.

407 **Segments.** We define the set of segments in several parts:

$$\begin{aligned}
 \text{chooseOr}_{i,j}^{\text{false}} &:= \{(q_{i,j}, r_{i,j}), (s_{i,j}, u_{i,j})\}, \\
 \text{chooseOr}_{i,j}^{\text{true}} &:= \{(m_{i,j}, s_{i,j}), (o_{i,j}, u_{i,j}), (t_{i,j}, v_{i,j})\}, \\
 \text{orMoveVariable}_{i,j} &:= \{(l_{i,j}, n_{i,j}), (n_{i,j}, p_{i,j})\}.
 \end{aligned}$$

408 Finally all segments on OR-gadget are defined as:

$$\text{segmentsOr}_{i,j} := \text{chooseOr}_{i,j}^{\text{false}} \cup \text{chooseOr}_{i,j}^{\text{true}} \cup \text{orMoveVariable}_{i,j}$$

409 **Lemma 3.5.** For any $1 \leq i \leq n, j \in \{0, 1\}$ and $x \in \{l_{i,j}, p_{i,j}\}$, points in $\text{pointsOr}_{i,j} - \{x\} \cup \{v_{i,j}\}$
 410 can be covered with 4 segments from $\text{segmentsOr}_{i,j}$.

411 *Proof.* We can do this using one segment from $\text{orMoveVariable}_{i,j}$, the one that does not cover
 412 x , and all segments from $\text{chooseOr}_{i,j}^{\text{true}}$. \square

413 **Lemma 3.6.** For any $1 \leq i \leq n, j \in \{0,1\}$, points in $\text{pointsOr}_{i,j}$ can be covered with 4
 414 segments from $\text{segmentsOr}_{i,j}$.

415 *Proof.* We can do this using segments from $\text{orMoveVariable}_{i,j} \cup \text{chooseOr}_{i,j}^{\text{false}}$. \square

416 3.3.3. CLAUSE-gadget

417 A CLAUSE-gadget is responsible for determining whether variable values assigned in variable
 418 gadgets satisfy the corresponding clause in the input formula ϕ . It has a minimum solution
 419 of size w if and only if the clause is satisfied, i.e. at least one of the respective variables is
 420 assigned the correct value. Otherwise, its minimum solution has size $w + 1$. In this way, by
 421 analyzing the size of the minimum solution to the entire constructed instance, we will be able
 422 to tell how many clauses it is possible to satisfy in an optimum solution to ϕ .



Figure 3.3: **CLAUSE-gadget for a clause $a \vee b \vee \neg c$.** Every green rectangle is an OR-gadget. y -coordinates of $x_{i,0}$, $y_{i,0}$ and $z_{i,0}$ depend on the variables in the i -th clause. Grey segments corresponds to the values of variables satisfying the i -th clause.

423 **Points.** First, we define auxiliary functions for literals. For a literal w , let $\text{idx}(w)$ be the
 424 index of the variable in w , and $\text{neg}(w)$ be the Boolean value (0 or 1) whether the variable is
 425 negated in w or not.

$$\begin{aligned} \text{idx}(w) &:= i \text{ when } w = x_i \\ \text{neg}(w) &:= \begin{cases} 0 & \text{if } w = x_i \\ 1 & \text{if } w = \neg x_i \end{cases} \end{aligned}$$

Let us assume that clause $C_i = a \vee b \vee c$ for any literals a, b, c . Then, we define points in the gadget as:

$$\begin{aligned} x_{i,0} &:= (20i, 4 \cdot \text{id}\mathbf{x}(a) + 2 \cdot \text{neg}(c)), & x_{i,1} &:= (20i, 4(n+1)), \\ y_{i,0} &:= (20i+1, 4 \cdot \text{id}\mathbf{x}(b) + 2 \cdot \text{neg}(b)), & y_{i,1} &:= (20i+1, 4(n+1)+4), \\ z_{i,0} &:= (20i+2, 4 \cdot \text{id}\mathbf{x}(c) + 2 \cdot \text{neg}(c)), & z_{i,1} &:= (20i+2, 4(n+1)+6). \end{aligned}$$

We are now ready to define the set of points in a clause gadget:

$$\begin{aligned} \text{moveVariable}_i &:= \{x_{i,j} : j \in \{0,1\}\} \cup \{y_{i,j} : j \in \{0,1\}\} \cup \{z_{i,j} : j \in \{0,1\}\}, \\ \text{pointsClause}_i &:= \text{moveVariable}_i \cup \text{pointsOr}_{i,0} \cup \text{pointsOr}_{i,1} \cup \{v_{i,1}\}. \end{aligned}$$

Note that these two points are equal: $v_{i,0} = l_{i,1}$. This translates to the fact that the output of the first OR-gadget is an input to the second OR-gadget. This creates an *or* of 3 boolean values.

Segments. We also define segments for the clause gadget as below:

$$\begin{aligned} \text{segmentsClause}_i &:= \{(x_{i,0}, x_{i,1}), (y_{i,0}, y_{i,1}), (z_{i,0}, z_{i,1}), (x_{i,1}, l_{i,0}), (y_{i,1}, p_{i,0}), (z_{i,1}, p_{i,1}), \} \cup \\ &\cup \text{segmentsOr}_{i,0} \cup \text{segmentsOr}_{i,1}. \end{aligned}$$

The CLAUSE-gadgets consist of two OR-gadgets. Ideally, we would place the i -th CLAUSE-gadget close to the $\mathbf{xTrueSegment}_{j_1}$ or $\mathbf{xFalseSegment}_{j_1}$ segments corresponding to the literals that occur in the i -th clause. It would be inconvenient to position them there, because between these segments there may be additional $\mathbf{xTrueSegment}_{j_2}$ or $\mathbf{xFalseSegment}_{j_2}$ segments corresponding to the other literals.

Instead, we use simple auxiliary gadgets to *transfer* whether the segment is in a solution, i.e. segments $(x_{i,0}, x_{i,1}), (y_{i,0}, y_{i,1}), (z_{i,0}, z_{i,1})$. Each transfer gadget consists of two segments $(x_{i,0}, x_{i,1}), (x_{i,1}, a)$. These are the only segments that can cover $x_{i,1}$. We place $x_{i,0}$ on a segment that we want to transfer (i.e. segment responsible for choosing the variable value satisfying the corresponding literal). If in some solution $x_{i,0}$ is already covered by this segment, then we can cover $x_{i,1}$ by $(x_{i,1}, a)$, thus also covering a . If $x_{i,0}$ is not covered by this segment, then the only way to cover $x_{i,0}$ is to use segment $(x_{i,0}, x_{i,1})$. Intuitively, in any optimal solution the two segments *transfer* the state of whether $x_{i,0}$ is covered onto whether a is covered. Therefore, the number of segments in the optimal solution is increased by one, and we get a point a that was effectively placed on some segment s , but it can be placed anywhere in the plane instead, consequently simplifying the construction.

Lemma 3.7. *For any $1 \leq i \leq n$ and $a \in \{x_{i,0}, y_{i,0}, z_{i,0}\}$, there is a set $\text{solClause}_i^{\text{true},a} \subseteq \text{segmentsClause}_i$ with $|\text{solClause}_i^{\text{true},a}| = 11$ that covers all points in $\text{pointsClause}_i - \{a\}$.*

Proof. For $a = x_{i,0}$ (analogous proof for $y_{i,0}$): First we use Lemma 3.5 twice with excluded $x = l_{i,0}$ and $x = l_{i,1} = v_{i,0}$, resulting with 8 segments in $\text{chooseOr}_{i,0}^{\text{true}} \cup \text{chooseOr}_{i,1}^{\text{true}}$ which cover all required points apart from $x_{i,1}, y_{i,0}, y_{i,1}, z_{i,0}, z_{i,1}, l_{i,0}$. We cover those using additional 3 segments: $\{(x_{i,1}, l_{i,0}), (y_{i,0}, y_{i,1}), (z_{i,0}, z_{i,1})\}$.

For $a = z_{0,i}$: Using Lemma 3.6 and Lemma 3.5 with $x = p_{i,1}$, we obtain 8 segments in $\text{chooseOr}_{i,0}^{\text{false}} \cup \text{chooseOr}_{i,1}^{\text{true}}$ which cover all required points apart from $x_{i,0}, x_{i,1}, y_{i,0}, y_{i,1}, z_{i,1}, p_{i,1}$. We cover those using additional 3 segments: $\{(x_{i,0}, x_{i,1}), (y_{i,0}, y_{i,1}), (z_{i,1}, p_{i,1})\}$. \square

Lemma 3.8. *For any $1 \leq i \leq n$ there is a set $\text{solClause}_i^{\text{false}} \subseteq \text{segmentsClause}_i$ with $|\text{solClause}_i^{\text{false}}| = 12$ that covers all points in pointsClause_i .*

Proof. Using Lemma 3.6 twice we can cover $\text{pointsOr}_{i,0}$ and $\text{pointsOr}_{i,1}$ with 8 segments. To cover the remaining points we additionally use: $\{(x_{i,0}, x_{i,1}), (y_{i,0}, y_{i,1}), (z_{i,0}, z_{i,1}), (t_{i,1}, v_{i,1})\}$.

□

Lemma 3.9. *For any $1 \leq i \leq n$:*

(1) *points in pointsClause_i can not be covered using any subset of segments from segmentsClause_i of size smaller than 12;*

(2) *points in $\text{pointsClause}_i - \{x_{i,0}, y_{i,0}, z_{i,0}\}$ can not be covered using any subset of segments from segmentsClause_i of size smaller than 11.*

Proof of (1). No segment in segmentsClause_i covers more than 1 point from

$$\{x_{i,0}, y_{i,0}, z_{i,0}, l_{i,0}, p_{i,0}, q_{i,0}, u_{i,0}, v_{i,0} = l_{i,1}, p_{i,1}, q_{i,1}, u_{i,1}, v_{i,1}\}.$$

Therefore we need to use at least 12 segments.

□

Proof of (2). We can define disjoint sets X, Y, Z such that

$$X \cup Y \cup Z \subseteq \text{pointsClause}_i - \{x_{i,0}, y_{i,0}, z_{i,0}\}$$

and there are no segments in segmentsClause_i covering points from different sets. And we prove a lower bound for each of these sets. First, let:

$$X := \{x_{i,1}, y_{i,1}, z_{i,1}\}.$$

No two points in X can be covered with one segment of segmentsClause_i , so it must be covered with 3 different segments. Next we define other sets:

$$Y := \text{pointsOr}_{i,0} - \{l_{i,0}, p_{i,0}\},$$

$$Z := \text{pointsOr}_{i,1} - \{l_{i,1}, p_{i,1}\}.$$

For both Y and Z we can check all of the subsets of 3 segments of segmentsClause_i to conclude that none of them cover the considered, so both Y and Z have to be covered with disjoint sets of 4 segments each.

Therefore, $\text{pointsClause}_i - \{x_{i,0}, y_{i,0}, z_{i,0}\}$ must be covered with at least $3 + 4 + 4 = 11$ segments from segmentsClause_i .

□

3.3.4. Summary

Finally we define the set of points and segments for the constructed instance:

$$\mathcal{C} := \bigcup_{1 \leq i \leq n} \text{pointsVariable}_i \cup \text{pointsClause}_i,$$

$$\mathcal{P} := \bigcup_{1 \leq i \leq n} \text{segmentsVariable}_i \cup \text{segmentsClause}_i.$$

Lemma 3.10 (Robustness to $\frac{1}{2}$ -extension). *For every segment $s \in \mathcal{P}$, s and $s^{+\frac{1}{2}}$ cover the same points from \mathcal{C} .*



Figure 3.4: **Scheme of the whole construction.**
General layout of VARIABLE-gadgets and CLAUSE-gadgets and how they interact with each other.

In order to prove this lemma we will define a claim about bounding rectangle for every gadget, which fits both segments and points from this gadget and after $\frac{1}{2}$ -extension does not cover any points from outside of this rectangle. Checking that the property from the above lemma holds for points and segments within the same gadget can be easily done using the figures above as references.

Note that the lemmas stated below also encapsulate the interactions between the gadgets, which are also mentioned in the helper lemmas above, and prove that gadgets are independent otherwise.

First let us define points to cover inside of rectangle R as:

$$\text{points}(R) := \text{points from } \mathcal{C} \text{ that lie in rectangle } R.$$

Claim 3.1. For any $1 \leq i \leq n$, $\text{segmentsVariable}_i$ fit in rectangle defined by points a_i and h_i from *VARIABLE-gadget*: $R := [-3L, L] \times [4i, 4i + 2]$.

- (1) The only points in R are pointsVariable_i and $x_{j,0}, y_{j,0}$ or $z_{j,0}$ points from *CLAUSE-gadget* for some clause c_j :

$$\text{pointsVariable}_i \subseteq \text{points}(R) \subseteq \text{pointsVariable}_i \cup \{x_{j,0}, y_{j,0}, z_{j,0} : 1 \leq j \leq n\}.$$

- (2) R covers the same points from \mathcal{C} before and after $\frac{1}{2}$ -extension, i.e. $\text{points}(R) = \text{points}(R^{+\delta})$.
- (3) All segments $\text{segmentsVariable}_i$ fit fully inside of R .

Claim 3.2. For any $1 \leq i \leq n$ and $j \in \{0, 1\}$, points from *OR-gadget* point $\text{pointsOr}_{i,j}$ and segments $\text{segmentsOr}_{i,j} - \{(t_{i,j}, v_{i,j})\}$ fit in rectangle defined as:

$$R := [x, x + 2] \times [y, y + 4], \text{ where } x = 20i + 3j + 3, y = 4(n + 1) + 2j.$$

- (1) R covers only $\text{pointsOr}_{i,j}$, i.e. $\text{points}(R) = \text{pointsOr}_{i,j}$.
- (2) R covers the same points from \mathcal{C} before and after $\frac{1}{2}$ -extension, i.e. $\text{points}(R) = \text{points}(R^{+\delta})$.
- (3) All segments $\text{segmentsOr}_{i,j} - \{(t_{i,j}, v_{i,j})\}$ fit fully inside of R .

Claim 3.3. For any $1 \leq i \leq n$, segmentsClause_i and pointsClause_i fit in rectangle:

$$R := [20i, 20i + 8] \times [0, 6(n + 1)].$$

- (1) R covers only pointsClause_i , i.e. $\text{points}(R) = \text{pointsClause}_i$.
- (2) R covers the same points from \mathcal{C} before and after $\frac{1}{2}$ -extension, i.e. $\text{points}(R) = \text{points}(R^{+\delta})$.
- (3) All segments segmentsClause_i fit fully inside of R .

Proof of Lemma 3.10. We will show that every gadget fits in their boundary rectangles with no interactions with other gadgets (even after $\frac{1}{2}$ -extensions) other than the ones described in lemmas described earlier.

To prove that points $x_{j,0}, y_{j,0}$ or $z_{j,0}$ are not covered by any segment from *VARIABLE* segments (TODO add symbol) after δ -extension: i -th *VARIABLE-gadget*'s points (TODO add symbol) fit in rectangle (TODO def of rectangle). Segments from (TODO add symbol) other than (TODO add true and false segments) also fit in this segment After δ -extension no other point from \mathcal{C} is covered by this rectangle.

Using claims 3.2 and 3.3. None of the segments in clause cover any new point after δ -ext.

□

3.4. Proof that the construction is correct and sound

In order to prove Lemma 3.1 we introduce several auxiliary lemmas proving properties of the construction described in the previous section.

Consider an instance S of MAX-(3,3)-SAT of size n with optimum solution satisfying k clauses. Let us construct an instance $(\mathcal{C}, \mathcal{P})$ of geometric set cover as described in Section 3.3 for the instance S of MAX-(3,3)-SAT.

Lemma 3.11. *The instance $(\mathcal{C}, \mathcal{P})$ of geometric set cover admits a solution of size $15n - k$.*

Proof. Let the clauses in S be c_1, c_2, \dots, c_n and the variables be x_1, x_2, \dots, x_n . Let the variable assignment in the optimum solution to S be $\phi : \{x_1, x_2, \dots, x_n\} \rightarrow \{\text{true}, \text{false}\}$.

We cover every VARIABLE-gadget with solution described in Lemma 3.2, where in the i -th gadget we choose the set of segments corresponding to the value of $\phi(x_i)$.

For every clause that is satisfied, say c_i , let us name the variable that is **true** in it as x_i and the point corresponding to x_i in **pointsClause_i** as a . Points in **pointsClause_i** are covered with set **solClause_i^{true,a}** described in Lemma 3.7. For every clause that is not satisfied, say c_j , points in **pointsClause_j** are covered with set **solClause_j^{false}** described in Lemma 3.8.

Formally, we define sets responsible for choosing variable assignment and satisfying clauses, R_i and C_i respectively, as following:

$$\begin{aligned} R_i &:= \begin{cases} \text{chooseVariable}_i^{\text{true}} & \text{if } \phi(x_i) = \text{true} \\ \text{chooseVariable}_i^{\text{false}} & \text{if } \phi(x_i) = \text{false} \end{cases} \\ C_i &:= \begin{cases} \text{solClause}_i^{\text{true},a} & \text{if } c_i \text{ satisfied by the literal corresponding to point } a \\ \text{solClause}_i^{\text{false}} & \text{if } c_i \text{ not satisfied} \end{cases} \\ \mathcal{R} &:= \bigcup_{i=1}^n \{R_i \cup C_i : 1 \leq i \leq n\}. \end{aligned}$$

This set covers all the points from \mathcal{C} , because the sets R_i, C_i individually cover their corresponding gadgets, as proved in the respective lemmas.

All of these sets are disjoint, so the size of the obtained solution is:

$$|\mathcal{R}| = \sum_{i=1}^n R_i + \sum_{i=1}^n C_i = 3n + 11k + 12(n - k) = 15n - k. \quad \square$$

Lemma 3.12. *Suppose we have a solution \mathcal{R} of the instance $(\mathcal{C}, \mathcal{P})$ of geometric set cover. Then there exists a solution \mathcal{R}' such that $|\mathcal{R}'| \leq |\mathcal{R}|$ and \mathcal{R}' contains at most one of the segments xTrueSegment_i and xFalseSegment_i from each VARIABLE-gadget.*

Proof. Assume that we have $\{\text{xTrueSegment}_i, \text{xFalseSegment}_i\} \subseteq \mathcal{R}$ for some i . We will show how to modify \mathcal{R} into \mathcal{R}' , such that the number of such i decreases, while \mathcal{R}' is still a valid solution to $(\mathcal{C}, \mathcal{P})$, and $|\mathcal{R}'| \leq |\mathcal{R}|$. Then, by repeating this procedure, we can eventually construct a solution satisfying the property from the Lemma.

To construct \mathcal{R}' , we first remove from \mathcal{R} all segments belonging to **segmentsVariable_i**. Recall that the i -th VARIABLE-gadget corresponds to variable x_i in S . As every variable in S is used in exactly 3 clauses, then one literal x_i or $\neg x_i$ must appear in at least 2 clauses. If that literal is x_i , then we add to the constructed solution all segments from **chooseVariable_i^{true}**, otherwise we add all segments from **chooseVariable_i^{false}**.

Now, there exists at most one CLAUSE-gadget which needs adjustment to make \mathcal{R}' valid; assuming it is the j -th clause, then one of the points $x_{j,0}, y_{j,0}$ or $z_{j,0}$ for this CLAUSE-gadget might be not covered, say $y_{j,0}$. We amend the solution by adding $(y_{j,0}, y_{j,1})$ to \mathcal{R}' .

By Lemma 3.4 we know that \mathcal{R} used at least 4 segments from $\text{segmentsVariable}_i$. Therefore, we removed at least 4 segments and added at most 4 segments, so $|\mathcal{R}'| \leq |\mathcal{R}|$. \square

Lemma 3.13. *Suppose we have a solution \mathcal{R} of the instance $(\mathcal{C}, \mathcal{P})$ of geometric set cover. Then there exists a solution to S that satisfies at least $15n - |\mathcal{R}|$ clauses.*

Proof. Let the clauses in S be c_1, c_2, \dots, c_n and the variables be x_1, x_2, \dots, x_n . Given a solution \mathcal{R} of the instance $(\mathcal{C}, \mathcal{P})$ of geometric set cover, we use Lemma 3.12 to modify \mathcal{R} so that for any i , \mathcal{R} contains at most one of xTrueSegment_i and xFalseSegment_i ; this may decrease the size of \mathcal{R} , but that does not matter in the subsequent construction. To simplify notation, in the remainder of this proof we use \mathcal{R} to refer to the modified solution.

Given \mathcal{R} , we construct a solution to S by defining an assignment of variables:

$$\phi : \{x_1, x_2, \dots, x_n\} \rightarrow \{\text{true}, \text{false}\}$$

that satisfies at least $15n - |\mathcal{R}|$ clauses in S .

Definition of ϕ . Recall that due to Lemma 3.12, \mathcal{R} contains at most one of xTrueSegment_i and xFalseSegment_i .

We define the value $\phi(x_i)$ for the variable x_i as follows:

$$\phi(x_i) := \begin{cases} \text{true} & \text{if } \text{xTrueSegment}_i \in \mathcal{R}, \\ \text{false} & \text{otherwise} \end{cases}$$

Moreover, from Lemma 3.3 we get $|\text{segmentsVariable}_i \cap \mathcal{R}| \geq 3$ for every i .

Clauses satisfied with the chosen variable assignment. For a clause c_i , \mathcal{R} needs to use at least 11 segments to cover $\text{pointsClause}_i - \{x_{i,0}, y_{i,0}, z_{i,0}\}$ in the i -th CLAUSE-gadget (Lemma 3.9).

Moreover, if none of the points $\{x_{i,0}, y_{i,0}, z_{i,0}\}$ are covered by the segments from $\mathcal{R} \cap \text{segmentsVariable}_i$, then \mathcal{R} needs to cover pointsClause_i with at least 12 segments by Lemma 3.9.

Let a be the number of clauses c_i for which none of the points $x_{i,0}, y_{i,0}, z_{i,0}$ in pointsClause_i are covered by segments from $\mathcal{R} \cap \text{segmentsVariable}_j$ for any $1 \leq j \leq n$.

Consider a clause c_i for which at least one of the points $x_{i,0}, y_{i,0}, z_{i,0}$ in pointsClause_i is covered by segments from $\mathcal{R} \cap \text{segmentsVariable}_j$ for some $1 \leq j \leq n$. Denote this point as t and say it corresponds to literal q and variable x_j . Point t can be only covered in $\text{segmentsVariable}_j$ by a corresponding segment xTrueSegment_j or xFalseSegment_j (depending on whether the literal q is negated or not). From the definition of ϕ and the fact that one of this segment is in \mathcal{R} , we know that $\phi(j)$ has the value that evaluates q to be **true**. Therefore, clause c_i is satisfied.

Consequently, ϕ satisfies all but at most a clauses in S .

To conclude, given a solution \mathcal{R} to $(\mathcal{C}, \mathcal{P})$ we constructed a variable assignment ϕ that satisfies at least $n - a$ clauses of S . Finally, note that

$$|\mathcal{R}| \geq 3n + 11(n - a) + 12a = 3n + 11n + a = 14n + a,$$

hence

$$15n - |\mathcal{R}| \leq 15n - 14n - a = n - a.$$

Therefore, ϕ satisfies at least $15n - |\mathcal{R}|$ clauses of S . \square

Chapter 4

Fixed-parameter tractable algorithm for geometric set cover problem

In this chapter we show fixed-parameter tractable algorithms for the geometric set cover problem in two different settings. Section 4.1 shows a fixed-parameter tractable algorithm for geometric set cover with unweighted segments. The remainder of the chapter presents a fixed-parameter tractable algorithm for geometric set cover with weighted segments with δ -extension. We show an algorithm for the setting with δ -extension, because the original problem with weights is W[1]-hard, as we show in Chapter 5.

We start with a shared definition for this problem. We define *extreme points* for a set of collinear points.

Definition 4.1. For a set of collinear points C in the plane, **extreme points** of C are the endpoints of the smallest segment that covers all points from set C .

If C consists of one point or is empty, then there are 1 or 0 extreme points respectively.

4.1. Fixed-parameter tractable algorithm for unweighted segments

In this section we consider fixed-parameter tractable algorithms for unweighted geometric set cover with segments. The setting where segments are required to be axis-parallel (or limited to a constant number of directions) has a trivial FPT algorithm. We present an FPT algorithm for geometric set cover with unweighted segments, where segments are in arbitrary directions.

4.1.1. Axis-parallel segments

Theorem 4.1. (*FPT for segment cover with axis-parallel segments*). There exists an algorithm that given a family \mathcal{P} of axis-parallel segments, a set of points \mathcal{C} and a parameter k , runs in time $\mathcal{O}(2^k)$, and outputs a solution $\mathcal{R} \subseteq \mathcal{P}$ such that $|\mathcal{R}| \leq k$ and \mathcal{R} covers all points in \mathcal{C} , or determines that such a set \mathcal{R} does not exist.

Proof. We show an $\mathcal{O}(2^k)$ -time branching algorithm. In each step, the algorithm selects a point a which is not yet covered, branches to choose one of the two directions, and greedily chooses a segment a in that direction to cover. This proceeds until either all points are covered or k segments are chosen.

Let us take the point $a = (x_a, y_a)$ which is the smallest among points that are not yet covered in the lexicographic ordering of points in \mathbb{R}^2 . We need to cover a with some of the remaining segments.

Branch over the choice of one of the coordinates (x or y); without loss of generality, let us assume we chose x . Among the segments lying on line $x = x_a$, we greedily add to the solution the one that covers the most points. As a was the smallest in the lexicographical order, all points on the line $x = x_a$ have the y -coordinate larger than y_a . Therefore, if we denote the greedily chosen segment as s , then any other segment on the line $x = x_a$ that covers a can only cover a subset of points covered by s . Thus, greedily choosing s is optimal.

In each step of the algorithm we add one segment to the solution, thus the recursion can be stopped at depth k . If no branch finds a solution, then this means that a solution of size at most k does not exist. \square

Note that the same algorithm can be used for segments in d directions, where we branch over d choices of directions, and it runs in complexity $\mathcal{O}(d^k)$.

4.1.2. Segments in arbitrary directions

In this section we consider the setting where segments are not constrained to a constant number of directions. We present a fixed-parameter tractable algorithm, parameterized by the size of the solution.

Theorem 1.2. (FPT for SEGMENT SET COVER). *There exists an algorithm that given a family \mathcal{P} of segments (in any direction), a set of points \mathcal{C} and a parameter k , runs in time $k^{\mathcal{O}(k)}(|\mathcal{C}| \cdot |\mathcal{P}|)^2$, and outputs a solution $\mathcal{R} \subseteq \mathcal{P}$ such that $|\mathcal{R}| \leq k$ and \mathcal{R} covers all points in \mathcal{C} , or determines that such a set \mathcal{R} does not exist.*

We will need the following lemmas proving properties of any instance of the problem.

Lemma 4.1. *Given an instance $(\mathcal{P}, \mathcal{C})$ of the segment cover problem, without loss of generality we can assume that no segment covers a superset of what another segment covers. That is, for any distinct $A, B \in \mathcal{P}$, we have $A \cap \mathcal{C} \not\subseteq B \cap \mathcal{C}$ and $A \cap \mathcal{C} \not\supseteq B \cap \mathcal{C}$.*

Proof. Assume towards a contradiction that there is an instance $(\mathcal{P}, \mathcal{C})$, and two distinct subsets of \mathcal{P} , A, B , such that $A \cap \mathcal{C} \subseteq B \cap \mathcal{C}$.

We construct a set $\mathcal{P}' := \mathcal{P} - \{A\}$. We prove that for any solution \mathcal{R} of $(\mathcal{P}, \mathcal{C})$, we can construct a solution $\mathcal{R}' \subseteq \mathcal{P}'$, such that $|\mathcal{R}'| \leq |\mathcal{R}|$. Let us take any solution \mathcal{R} of $(\mathcal{P}, \mathcal{C})$. If $A \in \mathcal{R}$, then $\mathcal{R}' := \mathcal{R} \cup \{B\} - \{A\}$, otherwise $\mathcal{R}' := \mathcal{R}$. Let us consider the case when $A \in \mathcal{R}$, because the other case is trivial. Since $A \cap \mathcal{C} \subseteq B \cap \mathcal{C}$, then $\mathcal{R} \cup \{B\} - \{A\}$ covers any point from \mathcal{C} that was covered by \mathcal{R} . Also, $|\mathcal{R} \cup \{B\} - \{A\}| \leq |\mathcal{R}|$. \square

Lemma 4.2. *Given an instance $(\mathcal{P}, \mathcal{C})$ of the segment cover problem transformed by Lemma 4.1, if there exists a line L with at least $k + 1$ points on it, then there exists a subset $A \subseteq \mathcal{P}$, of size at most k , such that every solution \mathcal{R} with $|\mathcal{R}| \leq k$ satisfies $|A \cap \mathcal{R}| \geq 1$. Moreover, such a subset can be found in polynomial time.*

Proof. Let us enumerate the points from \mathcal{C} that lie on L as x_1, x_2, \dots, x_t in the order in which they appear on L . Our proposed set is defined as:

$$A := \{\text{segment collinear with } L \text{ that covers } x_i \text{ and does not cover } x_{i-1} : i \in \{1, \dots, k\}\},$$

where for $i = 1$ we just take a segment that covers x_1 . If such a segment does not exist for any point x as above, then x does not give rise to any segment in A .

644 We prove the lemma by contradiction. Let us assume that there exists a solution \mathcal{R} of
645 size at most k such that $\mathcal{R} \cap A = \emptyset$.

646 Let \mathcal{R}_L be the set of segments from \mathcal{R} that are collinear with L .

647 Every segment that is not collinear with L can cover at most one of the points that lie
648 on this line. Hence, if \mathcal{R}_L was empty, then \mathcal{R} would cover at most k points on line L , but L
649 had at least $k + 1$ different points from \mathcal{C} on it.

650 Therefore, we know that \mathcal{R}_L is not empty and $|\mathcal{R} - \mathcal{R}_L| \leq k - 1$. Segments from $\mathcal{R} - \mathcal{R}_L$
651 can cover at most $k - 1$ points among $\{x_1, x_2, \dots, x_k\}$, therefore at least one of these points
652 must be covered by segments from \mathcal{R}_L . We take the leftmost point from $\{x_1, x_2, \dots, x_k\}$ that
653 is covered in \mathcal{R}_L and name it a . After the transformation from Lemma 4.1, in \mathcal{R} there is only
654 one segment that starts in a and is collinear with L , therefore this segment must be in both
655 \mathcal{R} and A . This contradiction concludes the proof that $|A \cap \mathcal{R}| \geq 1$ for any solution \mathcal{R} of size
656 at most k . \square

657 We are now ready to prove Theorem 1.2.

658 *Proof of Theorem 1.2.* We will prove this theorem by presenting a branching algorithm that
659 works in desired complexity. It first branches over the choice of segments to cover the lines
660 with *many* points and then solves a small instance (where every line has at most k points) by
661 checking all possible solutions.

662 **Algorithm.** We present a recursive algorithm. Given an instance of the problem:

- 663 (1) Use Lemma 4.1 to remove some redundant segments from our instance.
- 664 (2) If there exists a line with at least $k + 1$ points from \mathcal{C} , we branch over the choice of
665 adding to the solution one of the at most k possible segments provided by Lemma 4.2;
666 name this segment s and name the set of points from \mathcal{C} that lie on s as S . By recursion,
667 we find a solution \mathcal{R} for the instance $(\mathcal{C} - S, \mathcal{P} - \{s\})$, and parameter $k - 1$. We return
668 $\mathcal{R} \cup \{s\}$. Note that if Lemma 4.2 returned \emptyset , then we respond NO.
- 669 (3) If every line has at most k points on it and $|\mathcal{C}| > k^2$, then answer NO.
- 670 (4) If $|\mathcal{C}| \leq k^2$, solve the problem by brute force: check all subsets of \mathcal{P} of size at most k .

671 **Correctness.** Lemma 4.2 proves that at least one segment that we branch over in (1)
672 must be present in every solution \mathcal{R} with $|\mathcal{R}| \leq k$. Therefore, the recursive call can find
673 a solution, provided there exists one.

674 In (2) the answer is no, because every line covers no more than k points from \mathcal{C} , which
675 implies the same about every segment from \mathcal{P} . Under this assumption we can cover only k^2
676 points with a solution of size k , which is less than $|\mathcal{C}|$.

677 Checking all possible solutions in (3) is trivially correct.

678 **Complexity.** In the leaves of the recursion we have $|\mathcal{C}| \leq k^2$, so $|\mathcal{P}| \leq k^4$, because
679 every segment can be uniquely identified by the two extreme points it covers (by Lemma 4.1).
680 Therefore, there are $\binom{k^4}{k}$ possible solutions to check, each can be checked in time $\mathcal{O}(k|\mathcal{C}|)$.
681 Thus, (3) takes time $k^{\mathcal{O}(k)}$.

682 In this branching algorithm our parameter k is decreased with every recursive call, so we
683 have at most k levels of recursion with branching over k possibilities. Candidates to branch
684 over can be found on each level in time $\mathcal{O}((|\mathcal{C}| \cdot |\mathcal{P}|)^{\mathcal{O}(1)})$.

Reduction from Lemma 4.1 can be implemented in time $\mathcal{O}((|\mathcal{C}| \cdot |\mathcal{P}|)^{\mathcal{O}(1)})$.

It follows that the overall complexity is $\mathcal{O}((|\mathcal{C}| \cdot |\mathcal{P}|)^{\mathcal{O}(1)} \cdot k^{\mathcal{O}(k)})$ \square

4.2. Fixed-parameter tractable algorithm for weighted segments with δ -extension

In this section we consider the geometric set cover problem for weighted segments relaxed with δ -extension. We show that this problem admits an FPT algorithm when parameterized by the size of the solution and δ . In the next chapter we show that the assumption about the problem being relaxed with δ -extension is necessary: we prove that geometric set cover problem for weighted segments (without extension) is W[1]-hard, which means there does not exist any FPT algorithm parameterized by solution size for it, assuming $\text{FPT} \neq \text{W}[1]$.

Theorem 1.3. (*FPT for WEIGHTED SEGMENT SET COVER with δ -extension*). *There exists an algorithm that given a family \mathcal{P} of n weighted segments (in any direction), a set of m points \mathcal{C} , and parameters k and $\delta > 0$, runs in time $f(k, \delta) \cdot (nm)^c$ for some computable function f and a constant c and outputs a set \mathcal{R} such that:*

- $\mathcal{R} \subseteq \mathcal{P}$,
- $|\mathcal{R}| \leq k$,
- $\mathcal{R}^{+\delta}$ covers all points in \mathcal{C} ,
- the weight of \mathcal{R} is not greater than the weight of an optimum solution of size at most k for this problem without δ -extension,

or determines that there is no set \mathcal{R} with $|\mathcal{R}| \leq k$ such that \mathcal{R} covers all points in \mathcal{C} .

To solve this problem we will introduce a lemma about choosing a *dense* subset of points. A dense subset of points for a set of collinear points C and parameters k and δ is a subset of C such that if we cover it with at most k segments, these segments after δ -extension will cover all of the points from C . We will prove that such set of size bounded by some function $f(k, \delta)$ always exists (Lemma 4.3). Later, Lemma 4.3 will allow us to find a kernel for our original problem.

Definition 4.2. For a set of collinear points C , a subset $A \subseteq C$ is (k, δ) -**dense** if for any set of segments R that covers A and such that $|R| \leq k$, it holds that $R^{+\delta}$ covers C .

Lemma 4.3. *For any set of collinear points C , $\delta > 0$ and $k \geq 1$, there exists a (k, δ) -dense set $A \subseteq C$ of size at most $(2 + \frac{2}{\delta})^k$. Moreover, there exists an algorithm that computes the (k, δ) -dense set in time $\mathcal{O}(|C| \cdot (2 + \frac{2}{\delta})^k)$.*

Proof. We prove this for a fixed δ by induction on k .

Inductive hypothesis. For any set of collinear points C , there exists a set A such that:

- A is subset of C ,
- A is (ℓ, δ) -dense for every $1 \leq \ell \leq k$,
- $|A| \leq (2 + \frac{2}{\delta})^k$,
- the extreme points of C are in A .

722 **Base case for $k = 1$.** It is sufficient that A consists of the extreme points of C .
723 If they are covered with one segment, it must be a segment that includes the extreme
724 points from C , so it covers the whole set C .
725 There are at most 2 extreme points in C and $2 < 2 + \frac{2}{\delta}$.

726 **Inductive step.** Assuming inductive hypothesis for any set of collinear points C and
727 for parameter k , we will prove it for $k + 1$.

728 Let s be the minimal segment that includes all points from C . That is, the extreme points
729 of C are endpoints of s .

730 We define $M = \lceil 1 + \frac{2}{\delta} \rceil$ subsegments of s by splitting s into M closed segments of equal
731 length. We name these segments v_i , note that $|v_i| = \frac{|s|}{M}$ for each $1 \leq i \leq M$.

732 Let C_i be the subset of C consisting of points lying on v_i .

733 Let t_i be the segment with endpoints being the extreme points of C_i . It might be a
734 degenerate segment if C_i consists of one point, or t_i might be empty if C_i is empty.

735 Figure 4.1 presents an example of such segments v_i and t_i .

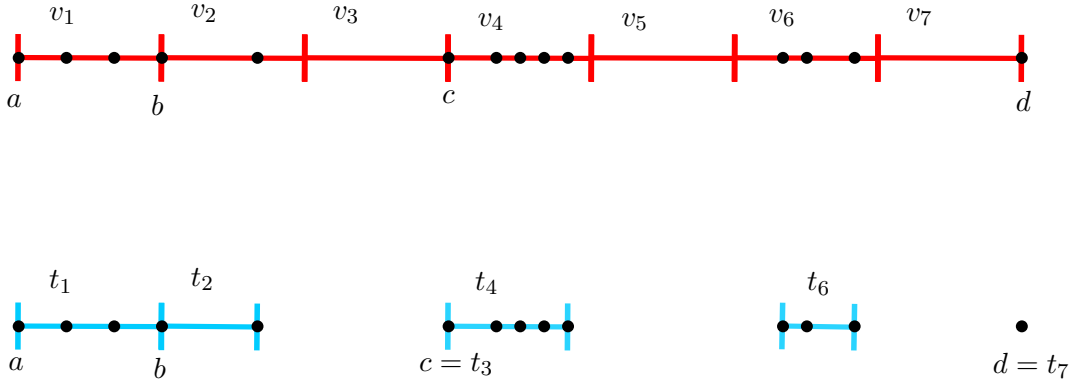


Figure 4.1: **Example of segments v_i and t_i .**

Example for $M = 7$ and some set of points (marked with black circles). The top panel shows segments v_i and the bottom panel shows segments t_i on the same set of points. a and b are the extreme points and therefore segment s ends at a and b . Red segments depict the split into M segments of equal length v_i . Blue segments depict the segments t_i . t_5 is an empty segment, because there are no points that lie on segment v_5 . Segments t_3 and t_7 are degenerated to one point – c and d , respectively. Segments t_1 and t_2 share one point b .

736 We use the inductive hypothesis to choose (k, δ) -dense sets A_i for sets C_i . Note that if
737 $|C_i| \leq 1$, then $A_i = C_i$ and it is still a (k, δ) -dense set for C_i .

738 Then we define $A = \bigcup_{i=1}^M A_i$. Thus A includes the extreme points of C , because they are
739 included in the sets A_1 and A_M .

The size of each A_i is at most $(2 + \frac{2}{\delta})^k$ from the inductive hypothesis, therefore size of A is at most:

$$M \left(2 + \frac{2}{\delta}\right)^k = \left\lceil 1 + \frac{2}{\delta} \right\rceil \cdot \left(2 + \frac{2}{\delta}\right)^k \leq \left(2 + \frac{2}{\delta}\right)^{k+1}.$$

740 **Proof that A is $(k+1, \delta)$ -dense for C .** Let us take any cover of A with $k+1$ segments
741 and call it \mathcal{R} .

742 For every segment t_i , if there exists a segment x in \mathcal{R} that is disjoint with t_i , then we have
743 a cover of A_i with at most k segments using $\mathcal{R} - \{x\}$. Since A_i is (k, δ) -dense for t_i and C_i ,
744 $(\mathcal{R} - \{x\})^{+\delta}$ covers C_i . So $\mathcal{R}^{+\delta}$ covers C_i as well.
745 If there exists a segment t_i for which a segment x as defined above does not exist, then
746 all $k + 1$ segments that cover A_i intersect t_i . An example of such segments is depicted in
747 Figure 4.2. Let us consider any such t_i . By the inductive hypothesis, the endpoints of s
748 are in A_1 and A_M respectively, so \mathcal{R} must cover them. For each endpoint of s , there exists
749 a segment that contains this endpoint and intersects t_i . Let us call these two segments y
750 and z . It follows that: $|y| + |z| + |t_i| \geq |s|$. Since $|t_i| \leq |v_i| = \frac{|s|}{M} \leq \frac{|s|}{1+\frac{2}{\delta}} = \frac{|s|\delta}{\delta+2}$, we have
751 $\max(|y|, |z|) \geq |s|(1 - \frac{\delta}{\delta+2})/2 = \frac{|s|}{\delta+2}$.



Figure 4.2: **Example of all $k + 1$ segments intersecting one segment t_i .**
Both panels show the same set \mathcal{C} (black circles), the same as in Figure 4.1. The top panel shows blue segments t_i for $M = 7$. The bottom panel shows green segments – solution \mathcal{R} of size 4. All segments from \mathcal{R} intersect t_4 . Segments z and y are named in the figure.

After δ -extension, the longer of these segments will expand at both ends by at least:

$$\max(|y|, |z|)\delta \geq \frac{|s|\delta}{\delta+2} = \frac{|s|}{1+\frac{2}{\delta}} \geq \frac{|s|}{M} = |v_i| \geq |t_i|.$$

752 Therefore, the longer of segments y and z will cover the whole segment t_i after δ -extension.
753 We conclude that $\mathcal{R}^{+\delta}$ covers C_i .
754 Since $C = \bigcup_{i=1}^M C_i$, it follows that $\mathcal{R}^{+\delta}$ covers C .

Algorithm. We can simulate the inductive proof presented above by a recursive algorithm with the following complexity:

$$O\left(|C| + \frac{1}{\delta}\right) + O\left(|C| \cdot \left(2 + \frac{2}{\delta}\right)^k\right).$$

755 □

756 Let us now formulate some claims about the properties for the problem parameterized
757 by the solution size. These properties provide bounds for different objects in the problem
758 instance, which help us to find a small kernel for the problem or conclude that the optimum
759 solution to this instance must be, in terms of size, above some threshold.

760 **Definition 4.3.** A line in the plane is **long** if there are at least $k + 1$ points from \mathcal{C} on it.

761 **Claim 4.1.** *If there are more than k different long lines, then \mathcal{C} can not be covered with k*
 762 *segments.*

763 *Proof.* We prove the claim by contradiction. Let us assume that we have at least $k+1$ different
 764 long lines in our instance of the problem and there is a solution \mathcal{R} of size at most k covering
 765 points \mathcal{C} .

766 Choose any long line L . Every segment from \mathcal{R} which is not collinear with L , covers at
 767 most one point that lies on L . L is long, so there are at least $k+1$ points from \mathcal{C} that lie on
 768 L . This implies that there must be a segment in \mathcal{R} that is collinear with L .

769 Since we have at least $k+1$ different long lines, there are at least $k+1$ segments in \mathcal{R}
 770 collinear with different lines. This contradicts with the assumption that $|\mathcal{R}| \leq k$. \square

771 **Claim 4.2.** *If there are more than k^2 points from \mathcal{C} that do not lie on any long line, then \mathcal{C}*
 772 *can not be covered with k segments.*

773 *Proof.* We prove the claim by contradiction. Let us assume that we have at least k^2+1 points
 774 from \mathcal{C} that do not lie on any long line, call this set A , and a solution \mathcal{R} of size at most k
 775 covering all points in \mathcal{C} .

776 Every segment s from \mathcal{R} covers at most k points from A . This is because if s covered at
 777 least $k+1$ points from A , then the line in the direction of s would be a long line and that
 778 contradicts the definition of A .

779 If every segment from \mathcal{R} covers at most k points from A and $|\mathcal{R}| \leq k$, then at most k^2
 780 points from A are covered by \mathcal{R} and that contradicts the fact that \mathcal{R} is a solution to the given
 781 geometric set cover instance. \square

782 We are now ready to give a proof of Theorem 1.3.

783 *Proof of Theorem 1.3.* Our goal is to either answer NO or to find a kernel $(\mathcal{C}', \mathcal{P}')$ of size
 784 bounded by $f(k)$ for some function f , such that:

- 785 • (*Property 1*) for every solution \mathcal{R} to $(\mathcal{C}, \mathcal{P})$ of size at most k , there exists a set $\mathcal{R}_1 \subseteq \mathcal{P}'$
 786 such that $|\mathcal{R}_1| \leq k$, the weight of \mathcal{R}_1 is not greater than the weight of \mathcal{R} , and \mathcal{R}_1 covers
 787 \mathcal{C}' ;
- 788 • (*Property 2*) for every set $\mathcal{R}_2 \subseteq \mathcal{P}'$ such that $|\mathcal{R}_2| \leq k$ and \mathcal{R}_2 covers all points in \mathcal{C}' ,
 789 $\mathcal{R}_2^{+\delta}$ covers all points in the original set \mathcal{C} .

790 If we found such sets $(\mathcal{C}', \mathcal{P}')$, using *Property 1* we know that an optimum solution of size
 791 at most k to $(\mathcal{C}', \mathcal{P}')$ has no greater weight than an optimum solution of size at most k to
 792 $(\mathcal{C}, \mathcal{P})$. Using *Property 2* we know that any solution to $(\mathcal{C}', \mathcal{P}')$ after δ -extension covers \mathcal{C} .

793 Therefore, finding such sets and solving the instance $(\mathcal{C}', \mathcal{P}')$ by iterating over all of the
 794 subsets of \mathcal{P}' of size at most k in desired complexity is sufficient to prove Theorem 1.3.

795 **Definition of \mathcal{C}' and \mathcal{P}' .** Let us name the number of different long lines as l . Applying
 796 Claims 4.1 and 4.2, if we have more than k different long lines or more than k^2 points from
 797 \mathcal{C} that do not lie on any long line, then we answer NO, because these lemmas prove that there
 798 is no solution of size at most k to this instance.

799 Otherwise, we can split \mathcal{C} into at most $k+1$ sets:

- 800 • D : points that do not lie on any long line, $|D| \leq k^2$;
- 801 • C_i for $1 \leq i \leq l$: points that lie on the i -th long line, $|C_i| > k$.

Note that sets C_i do not need to be disjoint.

Then, for every set C_i we can use Lemma 4.3 to obtain a (k, δ) -dense set A_i for C_i with $|A_i| \leq (2 + \frac{2}{\delta})^k$.

We define $\mathcal{C}' := D \cup (\bigcup A_i)$. \mathcal{C}' has size at most $k^2 + k(2 + \frac{2}{\delta})^k$. We define \mathcal{P}' as follows: for every pair of points \mathcal{C}' , we choose one segment from \mathcal{P} that has the lowest weight among segments that cover these points or decide that there is no segment that covers them. There are at most $|\mathcal{C}'|^2$ different segments in \mathcal{P}' , therefore both \mathcal{P}' and \mathcal{C}' have size bounded by $\mathcal{O}((k^2 + k(2 + \frac{2}{\delta})^k)^2)$.

Proof of Property 2. Firstly, we prove that for every set $\mathcal{R}_2 \subseteq \mathcal{P}'$ such that $|\mathcal{R}_2| \leq k$ and \mathcal{R}_2 covers points in \mathcal{C}' , $\mathcal{R}_2^{+\delta}$ covers points in the original instance \mathcal{C} .

Let us take such a set \mathcal{R}_2 .

\mathcal{C} is partitioned into several parts – sets D and C_i . Points from D are covered by \mathcal{R}_2 , because D is part of \mathcal{C}' . Each point from any A_i is covered, because A_i is a part of \mathcal{C}' ; A_i is a (k, δ) -dense set for C_i , therefore $\mathcal{R}_2^{+\delta}$ covers all points in C_i . Therefore, $\mathcal{R}_2^{+\delta}$ covers all points in \mathcal{C} .

Proof of Property 1. Secondly, we prove that for every solution \mathcal{R} to $(\mathcal{C}, \mathcal{P})$ of size at most k , there exists a set $\mathcal{R}_1 \subseteq \mathcal{P}'$ such that $|\mathcal{R}_1| \leq k$, the weight of \mathcal{R}_1 is not greater than the weight of \mathcal{R} and \mathcal{R}_1 covers \mathcal{C}' .

For every segment in \mathcal{R} , say s , let us look at the points from \mathcal{C}' that lie on s and call this set of points F . F is of course a set of collinear points. We can cover F with any segment that covers extreme points of F , because all other points lie on the segment between these points. Therefore, we can replace s with a segment s' that has lowest weight among the points that cover the extreme points of F . Such a segment belongs to \mathcal{P}' , because this is how it was defined. Segment s' has weight no greater than the weight of s , because s also covers F .

Therefore, we produced the set \mathcal{R}_1 that has size not greater than the size of \mathcal{R} (because some segments s can map to the same segment s'), weight not greater than \mathcal{R} , and it covers \mathcal{C}' .

Complexity We find a solution of $(\mathcal{C}', \mathcal{P}')$ by iterating over all the possible subsets of \mathcal{P}' . Finding sets \mathcal{P}' and \mathcal{C}' and then solving problem for kernel has overall complexity $(|\mathcal{P}| + |\mathcal{C}|)^{\mathcal{O}(1)} \mathcal{O}((2 + \frac{2}{\delta})^k) + \mathcal{O}((k^2 + k(2 + \frac{2}{\delta})^k)^k)$. \square

Chapter 5

W[1]-hardness for axis-parallel weighted segments

In this chapter we consider the geometric set cover problem with axis-parallel or right-diagonal weighted segments. In Theorem 1.4 below, we prove that this problem is W[1]-hard when parameterized by the size of the solution.

We believe that the below construction can be improved to only utilize the axis-parallel segments.

Theorem 1.4. *Consider the problem of covering a set \mathcal{C} of points by selecting at most k segments from a set of segments \mathcal{P} with non-negative weights $w : \mathcal{P} \rightarrow \mathbb{R}^+$ so that the weight of the cover is minimal. Then this problem is W[1]-hard when parameterized by k and assuming ETH, there is no algorithm for this problem with running time $f(k) \cdot (|\mathcal{C}| + |\mathcal{P}|)^{o(\sqrt{k})}$ for any computable function f . Moreover, this holds even if all segments in \mathcal{P} are axis-parallel or right-diagonal.*

5.1. Grid Tiling

In order to prove Theorem 1.4 we will show a reduction from a W[1]-hard problem: grid tiling. This problem was introduced in [Marx, 2007] (the author called it matrix tiling instead). It was originally described as an approximation problem, but W[1]-hardness follows directly from the theorems stated there. For a more contemporary description of this problem and a proof of W[1]-hardness, see Chapter 14 of [Cygan et al., 2015].

Definition 5.1. We define the **powerset** of a set A , denoted as $\text{Pow}(A)$, as the set of all subsets of A , i.e. $\text{Pow}(A) = \{B : B \subseteq A\}$.

Definition 5.2. In the **grid tiling** problem we are given integers n and k , and a function $f : \{1, \dots, k\} \times \{1, \dots, k\} \rightarrow \text{Pow}(\{1, \dots, n\} \times \{1, \dots, n\})$ specifying the set of allowed tiles for each cell of a $k \times k$ grid. The task is to decide whether there exist functions $x, y : \{1, \dots, k\} \rightarrow \{1, \dots, n\}$ that assign colors from $\{1, \dots, n\}$ to respectively columns and rows of the grid, so that $(x(i), y(j)) \in f(i, j)$ for all $i, j \in \{1, \dots, k\}$.

In short, in the grid tiling problem one needs to assign numbers to rows and columns in such a way that for every pair of a row and a column, the pair of colors assigned to the row and column belongs to the allowed set of tiles for this pair. The next theorem describes the complexity of this problem, which is W[1]-hard when parameterized by the size of the grid.

	$x(1) = 3$	$x(2) = 1$	$x(3) = 3$	$x(4) = 7$
$y(4) = 1$	$(\mathbf{2}, \mathbf{1}); (2, 2);$ $(\mathbf{3}, \mathbf{1}); (3, 9)$	$(1, 1); (3, 1)$	$(\mathbf{3}, \mathbf{1}); (7, 2)$	$(\mathbf{2}, \mathbf{1}); (\mathbf{7}, \mathbf{1})$
$y(3) = 1$	$(\mathbf{2}, \mathbf{1}); (\mathbf{3}, \mathbf{1});$ $(4, 2); (8, 2)$	$(1, 1); (1, 3)$	$(\mathbf{3}, \mathbf{1}); (4, 3)$	$(\mathbf{2}, \mathbf{2}); (\mathbf{7}, \mathbf{1})$
$y(2) = 6$	$(\mathbf{2}, \mathbf{6}); (\mathbf{3}, \mathbf{6})$	$(1, 2); (\mathbf{1}, \mathbf{6});$ $(2, 6)$	$(2, 6); (\mathbf{3}, \mathbf{6})$	$(\mathbf{2}, \mathbf{6}); (\mathbf{7}, \mathbf{6})$
$y(1) = 4$	$(\mathbf{2}, \mathbf{4}); (2, 6);$ $(\mathbf{3}, \mathbf{4}); (\mathbf{3}, \mathbf{9})$	$(1, 4); (\mathbf{1}, \mathbf{9})$	$(\mathbf{3}, \mathbf{4}); (\mathbf{3}, \mathbf{9})$	$(\mathbf{2}, \mathbf{9}); (\mathbf{7}, \mathbf{4})$

Figure 5.1: **Example of a grid tiling instance and its solution.**

In the first row and column of the table you can see the solution: functions x and y . The tiles used in this solution are marked in **bold**. If we instead chose the tiles marked in **blue** (whenever there is one, taking the tile marked in **bold** otherwise), then that corresponds to setting $x(1) = 2$, and would also form a correct solution. On the other hand, if we instead chose the tiles marked in **red** (as before), then this corresponds to setting $y(1) = 9$ and $x(4) = 2$ and that would **not** form a correct solution. Even though the first row is correct, the cell with coordinates $(3, 4)$ requires tile $(2, 1)$, not $(2, 2)$ (marked in **bold red**).

Theorem 5.1. [Marx, 2007] *Grid tiling is $W[1]$ -hard when parameterized by k and assuming ETH, there is no $f(k) \cdot n^{o(k)}$ -time algorithm solving the grid tiling problem for any computable function f .*

The remainder of this section is devoted to proving Theorem 1.4 by a reduction from a grid tiling problem instance with parameter k (number of rows in the grid) to a geometric set cover instance with parameter k^2 (size of solution). This reduction is described in Lemma 5.1. This proves the $W[1]$ -hardness of the geometric set cover problem, because if we could solve it with an FPT algorithm, then we could also solve the grid tiling problem (which we reduced to the geometric set cover). Therefore, geometric set cover with setting described in Theorem 1.4 is at least as hard as the grid tiling problem.

5.2. Statement of reduction

Let us denote an instance of grid tiling problem as (n, k, f) consisting of:

- the number of colors n ,
- the size of the grid k ,
- the function specifying the allowed tiles $f : \{1, \dots, k\} \times \{1, \dots, k\} \rightarrow \text{Pow}(\{1, \dots, n\} \times \{1, \dots, n\})$.

Let us also define constants:

$$\begin{aligned} \epsilon &:= \frac{1}{2k^2} \\ \delta &:= \frac{1}{4k^4} \\ W_{\text{hv}} &:= 2k^2(n^2 + 1) - 4k^2\epsilon - 4k(1 - \epsilon) \end{aligned}$$

which are going to be used when defining the weight of the constructed instance of geometric set cover with weighted segments.

Lemma 5.1. *Given an instance (n, k, f) of the grid tiling problem, we can construct an instance $(\mathcal{C}, \mathcal{P}, w, 3k^2 + 2k)$ of geometric set cover with weighted segments such that:*

- (1) if the answer to (n, k, f) is YES, then there exists a solution to $(\mathcal{C}, \mathcal{P}, w, 3k^2 + 2k)$ of weight at most $W_{\text{hv}} + k^2\delta$;*
- (2) if there exists a solution to $(\mathcal{C}, \mathcal{P}, w, 3k^2 + 2k)$ of weight at most $W_{\text{hv}} + k^2\delta$, then the answer to (n, k, f) is YES.*

First, let us prove Theorem 1.4 using Lemma 5.1.

Proof of Theorem 1.4. Let us take any instance (n, l, f) of the grid tiling problem. We prove the theorem by contradiction, therefore we assume that geometric set cover with weighted segments parameterized by solution size k admits a $g(k) \cdot n^{o(\sqrt{k})}$ -time algorithm for some computable function g .

Using Lemma 5.1 let us construct an instance I for (n, l, f) . Let us assume that the optimum solution of size at most k to the instance I has weight u . Using (2) we know that if $u \leq W_{\text{hv}} + k^2\delta$, then the answer to (n, l, f) is YES. If $u > W_{\text{hv}} + k^2\delta$, then using (1) we know that the answer to (n, l, f) must be NO.

Therefore if we could find the solution in time $g(k) \cdot n^{o(\sqrt{k})}$, then we could solve the grid tiling problem in time $g(l) \cdot n^{o(l)}$ by constructing an instance of the set cover with weighted segments, solving it for parameter $k = 3l^2 + 2l$ in time $n^{o(\sqrt{3l^2+2l})}$ and then answering based on the weight of the optimum solution. As $\mathcal{O}(n^{o(l)}) \subseteq \mathcal{O}(n^{o(\sqrt{3l^2+2l})})$, the existence of this algorithm contradicts Theorem 5.1. Hence such an algorithm can not exist. \square

We prove Lemma 5.1 in subsequent sections. First, we define a constructed instance I , later property (1) is proved by Lemma 5.2 and property (2) is proved by Lemma 5.6.

In the proof of Lemma 5.6 we do not use the assumption that the solution is bounded by the size, which the problem is parameterized by, $3k^2 + 2k$. If we had a permissive FPT algorithm that finds a solution of any size that still has weight no more than $W_{\text{hv}} + k^2\delta$, then we still would have a contradiction with grid tiling being W[1]-hard in proof of Theorem 1.4. Thus, this reduction proves that the problem is not only W[1]-hard, but assuming ETH there also does not exist permissive FPT algorithm for this problem. Formally we state this in the Theorem 5.2.

Theorem 5.2. (Permissive FPT does not exist). *Consider the problem of covering a set \mathcal{C} of points using segments from a set \mathcal{P} with non-negative weights $w : \mathcal{P} \rightarrow \mathbb{R}^+$ so that the weight of the cover is minimal. Let \mathcal{R}^k be the optimum solution to this problem of size at most k . The task is to find a solution \mathcal{R} of any size such that weight of \mathcal{R} is not greater than the weight of \mathcal{R}^k .*

Assuming ETH, there is no algorithm for this problem with running time $f(k) \cdot (|\mathcal{C}| + |\mathcal{P}|)^{o(\sqrt{k})}$ for any computable function f . Moreover, this holds even if all segments in \mathcal{P} are axis-parallel or right-diagonal.

5.3. Construction of the Geometric Set Cover instance

We construct an instance $(\mathcal{C}, \mathcal{P}, w, 3k^2 + 2k)$ of geometric set cover as follows.

First, let us choose any bijection $\text{order} : \{1, \dots, n^2\} \rightarrow \{1, \dots, n\} \times \{1, \dots, n\}$.

Define $\text{match}_v(i, j)$ and $\text{match}_h(i, j)$ as boolean functions denoting whether two points share x or y coordinate:

$\text{match}_v(i, j)$ is **true** \iff $\text{order}(i)$ and $\text{order}(j)$ have the same x coordinate,

$\text{match}_h(i, j)$ is **true** \iff $\text{order}(i)$ and $\text{order}(j)$ have the same y coordinate.

922 5.3.1. Points

For $1 \leq i, j \leq k$ and $1 \leq t \leq n^2$ define points:

$$h_{i,j,t} := (i \cdot (n^2 + 1) + t, j \cdot (n^2 + 1)),$$

$$v_{i,j,t} := (i \cdot (n^2 + 1), j \cdot (n^2 + 1) + t).$$

Let us define sets H and V as:

$$H := \{h_{i,j,t} : 1 \leq i, j \leq k, 1 \leq t \leq n^2\},$$

$$V := \{v_{i,j,t} : 1 \leq i, j \leq k, 1 \leq t \leq n^2\}.$$

Let us recall that $\epsilon = \frac{1}{2k^2}$. For a point $p = (x, y)$ we define points:

$$p^L := (x - \epsilon, y),$$

$$p^R := (x + \epsilon, y),$$

$$p^U := (x, y + \epsilon),$$

$$p^D := (x, y - \epsilon).$$

Then we define the point set as follows:

$$\mathcal{C} := H \cup \{p^L : p \in H\} \cup \{p^R : p \in H\} \cup V \cup \{p^U : p \in V\} \cup \{p^D : p \in V\}.$$

923 **Definition 5.3.** For every point $p \in H$, we name point p^L its **left guard** and point p^R its
924 **right guard**.

925 Similarly for every points $p \in V$, we name point p^D its **lower guard** and point p^U its
926 **upper guard**.

927 5.3.2. Segments

928 For $1 \leq i, j \leq k$ and $1 \leq t, t_1, t_2 \leq n^2$ define segments:

$$\text{hor}_{i,j,t_1,t_2} := (h_{i,j,t_1}^R, h_{i+1,j,t_2}^L),$$

$$\text{ver}_{i,j,t_1,t_2} := (v_{i,j,t_1}^U, v_{i,j+1,t_2}^D),$$

$$\text{horBeg}_{i,t} := (h_{1,i,1}^L, h_{1,i,t}^L),$$

$$\text{horEnd}_{i,t} := (h_{k,i,t}^R, h_{k,i,n^2}^R),$$

$$\text{verBeg}_{i,t} := (v_{i,1,1}^D, v_{i,1,t}^D),$$

$$\text{verEnd}_{i,t} := (v_{i,k,t}^U, v_{i,k,n^2}^U).$$

929 Next, we define sets of vertical and horizontal segments:

$$\begin{aligned} \text{HOR} &:= \{ \text{hor}_{i,j,t_1,t_2} : 1 \leq i < k, 1 \leq j \leq k, 1 \leq t_1, t_2 \leq n^2, \text{match}_h(t_1, t_2) \text{ holds} \} \\ &\cup \{ \text{horBeg}_{i,t} : 1 \leq i \leq k, 1 \leq t \leq n^2 \} \\ &\cup \{ \text{horEnd}_{i,t} : 1 \leq i \leq k, 1 \leq t \leq n^2 \}, \end{aligned}$$

930

$$\begin{aligned} \text{VER} &:= \{ \text{ver}_{i,j,t_1,t_2} : 1 \leq i \leq k, 1 \leq j < k, 1 \leq t_1, t_2 \leq n^2, \text{match}_v(t_1, t_2) \text{ holds} \} \\ &\cup \{ \text{verBeg}_{i,t} : 1 \leq i \leq k, 1 \leq t \leq n^2 \} \\ &\cup \{ \text{verEnd}_{i,t} : 1 \leq i \leq k, 1 \leq t \leq n^2 \}. \end{aligned}$$

931 An example is depicted in Figure 5.3.

Finally, we also define a set of right-diagonal segments:

$$\text{DIAG} := \{ (h_{i,j,t}, v_{i,j,t}) : 1 \leq i, j \leq k, 1 \leq t \leq n^2, \text{order}(t) \in f(i, j) \}.$$

932 An example of such segments is depicted in Figure 5.2.

933 Every segment in **DIAG** connects points $(i(n^2+1)+t, j \cdot (n^2+1))$ and $(i \cdot (n^2+1), j(n^2+1) + t)$
 934 for some $1 \leq i, j \leq k, 1 \leq t \leq n^2$. The line on which it lies can be described by linear equation
 935 $x + y = t + (i + j)(n^2 + 1)$, thus these segments are in fact right-diagonal.

936 The constructed segment set is defined as:

$$\mathcal{P} := \text{HOR} \cup \text{VER} \cup \text{DIAG}.$$

937 The weight of each segment in $\text{HOR} \cup \text{VER}$ is equal to its length, while every segment in
 938 **DIAG** has weight δ .

$$w(s) = \begin{cases} \text{length}(s) & \text{if } s \in \text{HOR} \cup \text{VER} \\ \delta & \text{if } s \in \text{DIAG} \end{cases}$$

939 5.4. Proof that reduction is correct

940 Now, we prove that the constructed instance of geometric set cover with weighted segments
 941 indeed gives a correct and sound reduction of the grid tiling problem. Lemma 5.2 proves that
 942 if a solution to the instance of the grid tiling instance exists, then there exists a solution with
 943 suitably bounded size and weight of the constructed instance of geometric set cover. Then
 944 Lemma 5.6 proves that if there is a solution to the geometric set cover instance with bounded
 945 weight, then there exists a solution to the original grid tiling instance.

946 **Lemma 5.2.** *If there exists a solution to the grid tiling instance $(f_{i,j})$, then there exists*
 947 *a solution to the instance $(\mathcal{C}, \mathcal{P}, w, 3k^2 + 2k)$ of geometric set cover with weight $W_{\text{hv}} + k^2\delta$.*

948 *Proof.* Suppose there exists a solution x, y of the instance $(f_{i,j})$ of the grid tiling problem.

949 We define the proposed solution $\mathcal{R} \subseteq \mathcal{P}$ of the instance of geometric set cover in three



Figure 5.2: **Vertices and segments in DIAG.**

This is an example of constructed points any $1 \leq i, j \leq k$. Points from H and V are marked in black, their guards are marked in blue. You can also see segments from DIAG with their weights (equal to δ).

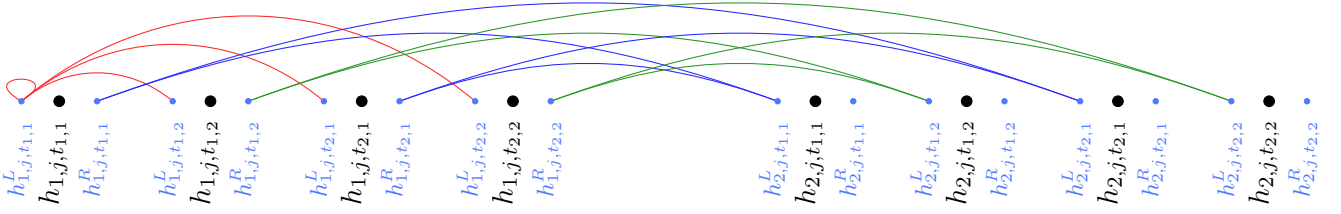


Figure 5.3: **Vertices and segments in HOR.**

This is an example for $n = 2$ and any $1 \leq j \leq k$. Points from H are marked in black, their guards are marked in light blue. $t_{i,j}$ is a notation that we use for $\text{order}^{-1}(i, j)$. Segments are represented as arcs between endpoints. You can see $\text{horBeg}_{j,t}$ segments in red. $\text{horBeg}_{j,1}$ is degenerated to a single point at $h_{1,1,t_{1,1}}^L$. Segments $\text{hor}_{i,j,t_{x_1,y},t_{x_2,y}}$ are marked in blue and green. Blue segments connect $t_{x_1,y}$ and $t_{x_2,y}$ such that they share y-coordinate equal to 1, for green segments it is equal to 2.

950 parts: $D \subseteq \text{DIAG}$, $A \subseteq \text{HOR}$ and $B \subseteq \text{VER}$:

$$\begin{aligned}
 D &:= \{(v_{i,j,t}, h_{i,j,t}) : 1 \leq i, j \leq k, t = \text{order}^{-1}(x(i), y(j))\}, \\
 A &:= \{\text{horBeg}_{i, \text{order}^{-1}(x(1), y(i))} : 1 \leq i \leq k\} \\
 &\quad \cup \{\text{horEnd}_{i, \text{order}^{-1}(x(k), y(i))} : 1 \leq i \leq k\} \\
 &\quad \cup \{\text{hor}_{i,j, \text{order}^{-1}(x(i), y(j)), \text{order}^{-1}(x(i+1), y(j))} : 1 \leq i < k, 1 \leq j \leq k\}, \\
 B &:= \{\text{verBeg}_{i, \text{order}^{-1}(x(i), y(1))} : 1 \leq i \leq k\} \\
 &\quad \cup \{\text{verEnd}_{i, \text{order}^{-1}(x(i), y(k))} : 1 \leq i \leq k\} \\
 &\quad \cup \{\text{ver}_{i,j, \text{order}^{-1}(x(i), y(j)), \text{order}^{-1}(x(i), y(j+1))} : 1 \leq i \leq k, 1 \leq j < k\}, \\
 \mathcal{R} &:= D \cup A \cup B.
 \end{aligned}$$

951 Since $\mathcal{C} = H \cup V$, we show that \mathcal{R} covers the whole set H ; the proof for V is analogous.

952 Fix any $1 \leq j \leq k$ and define $t_i := \text{order}^{-1}(x(i), y(j))$. The two leftmost segments in A
 953 for this j are $\text{horBeg}_{j,t_1} = (h_{1,j,1}^L, h_{1,j,t_1}^L)$ and $\text{hor}_{1,j,t_1,t_2} = (h_{1,j,t_1}^R, h_{2,j,t_2}^L)$. Therefore, points
 954 $h_{1,j,x}, h_{1,j,x}^L$ and $h_{1,j,x}^R$ for all $1 \leq x \leq n^2$ are covered by horBeg_{j,t_1} and hor_{1,j,t_1,t_2} , excluding
 955 point h_{1,j,t_1} .

956 Analogously for $2 \leq i \leq k-1$, the two consecutive segments $\text{hor}_{i-1,j,t_{i-1},t_i}$ and $\text{hor}_{i,j,t_i,t_{i+1}}$
 957 cover points $h_{i,j,x}, h_{i,j,x}^L$ and $h_{i,j,x}^R$ for all $1 \leq x \leq n^2$, excluding point h_{i,j,t_i} .

958 Finally $\text{hor}_{k-1,j,t_{k-1},t_k}$ and horEnd_{j,t_k} cover all points $h_{k,j,x}, h_{k,j,x}^L$ and $h_{k,j,x}^R$ for $1 \leq x \leq n^2$,
 959 excluding point h_{k,j,t_k} .

960 D covers all points h_{i,j,t_i} and v_{i,j,t_i} . As j was chosen arbitrarily, all points in H are covered.
 The size of this proposed solution is:

$$|\mathcal{R}| = |D| + |A| + |B| = k^2 + (k+1)k + (k+1)k = 3k^2 + 2k.$$

961 Then, we need to compute the total weight of the solution \mathcal{R} . First, we compute the sum
 962 of weights of segments in A . Fix $1 \leq j \leq k$ and consider segments collinear with the j -th
 963 horizontal line. All points $h_{i,j,t}, h_{i,j,t}^L$ and $h_{i,j,t}^R$ for every $1 \leq i \leq k$ and $1 \leq t \leq n^2$ are covered
 964 by A excluding points $h_{i,j, \text{order}^{-1}(x(i), y(j))}$. Every such point leaves a gap of length 2ϵ between
 965 $h_{i,j, \text{order}^{-1}(x(i), y(j))}^L$ and $h_{i,j, \text{order}^{-1}(x(i), y(j))}^R$. Therefore, the total weight of segments in A that
 966 lie on the line in question equals the length of the segment $(h_{1,1,1}^L, h_{i,k,n^2}^R)$ minus $2\epsilon k$, which is

967 $k(n^2 + 1) - 2(1 - \epsilon) - 2k\epsilon$. We need to multiply that by k , as we consider all possible values
 968 of j .

969 Computation for vertical segments is analogous and yields the same result. Every segment
 970 in D has weight δ , therefore the sum of all weights is equal to:

$$2k(k(n^2 + 1) - 2(1 - \epsilon) - 2k\epsilon) + k^2\delta = W_{\text{hv}} + k^2\delta. \quad \square$$

971 Now we present a few additional properties of the constructed instance of the geometric
 972 set cover that help us to prove Lemma 5.6.

973 **Claim 5.1.** *In any solution to the instance $(\mathcal{C}, \mathcal{P}, w, 3k^2 + 2k)$:*

- 974 • *the left and right guards of points in H (points in $\{p^L : p \in H\} \cup \{p^R : p \in H\}$) have*
 975 *to be covered with segments from HOR,*
- 976 • *the lower and upper guards of points in V (points in $\{p^D : p \in V\} \cup \{p^U : p \in V\}$) have*
 977 *to be covered with segments from VER.*

978 *Proof.* We prove the claim for the points from H as the proof for points from V is analogous.
 979 Every segment in VER is vertical and has x-coordinate equal to $i(n^2 + 1)$ for some $1 \leq i \leq k$,
 980 so they all have different x-coordinate than any left or right guard of points in H .

981 For every point x which is a left or right guard of a point in H , there are kn^2 segments
 982 from DIAG that intersect with the horizontal line that goes through x . All of these segments
 983 intersect with this line in points from set H , therefore none of them covers any of the guards.

984 Therefore none of the segments from VER or DIAG covers any of the guards of the points
 985 in H . \square

986 **Claim 5.2.** *For any $1 \leq i, j \leq n$ and any solution to the instance $(\mathcal{C}, \mathcal{P}, w, 3k^2 + 2k)$, all*
 987 *but at most one point $h_{i,j,t}$ and at most one point $v_{i,j,t}$ for $1 \leq t \leq n^2$ must be covered with*
 988 *segments from HOR or VER.*

989 *Proof.* We prove the claim for horizontal segments, as the proof for vertical segments is anal-
 990 ogous.

991 We prove this by contradiction. Assume that we have two points $h_{i,j,t_1}, h_{i,j,t_2}, 1 \leq t_1 <$
 992 $t_2 \leq n^2$, such that they are not covered with segments from HOR.

993 Point h_{i,j,t_1}^R has to be covered with a segment from HOR by Claim 5.1. Every segment in
 994 HOR covering h_{i,j,t_1}^R , but not h_{i,j,t_1} must start at h_{i,j,t_1}^R and all such segments cover also h_{i,j,t_2} .
 995 This contradicts the assumption, which concludes the proof. \square

996 **Lemma 5.3.** *For every solution \mathcal{R} to the instance $(\mathcal{C}, \mathcal{P}, w, 3k^2 + 2k)$, the sum of weights of*
 997 *segments chosen from sets HOR and VER is at least W_{hv} .*

998 *Proof.* Let us fix $1 \leq i \leq k$.

999 We provide a lower bound for the sum of lengths of vertical segments from $\mathcal{R} \cap \text{VER}$. This
 1000 bound is the same for each i and is the same for horizontal lines, thus we need to multiply
 1001 such a bound by $2k$.

(1) The total length between $v_{i,1,1}^D$ and v_{i,k,n^2}^U is:

$$(k(n^2 + 1) + n^2 + \epsilon) - ((n^2 + 1) + 1 - \epsilon) = k(n^2 + 1) - 2(1 - \epsilon).$$

1002 (2) For every $1 \leq j \leq k$ there exists at most one $1 \leq t \leq n^2$ such that $v_{i,j,t}$ is not covered
 1003 by segments from **VER** (Claim 5.2). Its guards (see Definition 5.3) $v_{i,j,t}^U$ and $v_{i,j,t}^D$ have
 1004 to be covered in **VER** (Claim 5.1). Therefore, at most k spaces of length 2ϵ can be left
 1005 not covered by segments from **VER** between $v_{i,1,1}^D$ and v_{i,k,n^2}^U .

The sum of these lower bounds for vertical and horizontal lines is:

$$2k(k(n^2 + 1) - 2k\epsilon - 2(1 - \epsilon)) = 2k^2(n^2 + 1) - 4k^2\epsilon - 4k(1 - \epsilon) = W_{\text{hv}}. \quad \square$$

1006 **Lemma 5.4.** *Let \mathcal{R} be a solution to a constructed instance $(\mathcal{C}, \mathcal{P}, w, 3k^2 + 2k)$ with weight at*
 1007 *most $W_{\text{hv}} + k^2\delta$. Then for every $1 \leq i, j \leq k$ there exists $1 \leq t \leq n^2$ such that:*

- 1008 (1) $v_{i,j,t}, h_{i,j,t}$ are not covered by segments from **VER** or **HOR**;
- 1009 (2) segment $(v_{i,j,t}, h_{i,j,t})$ is in solution \mathcal{R} ;
- 1010 (3) $\text{order}(t) \in f(i, j)$, that is, $\text{order}(t)$ is an allowed tile for (i, j) ;
- 1011 (4) for every $1 \leq s \leq n^2$, $s \neq t$, $v_{i,j,s}$ is covered in **VER**;
- 1012 (5) for every $1 \leq s \leq n^2$, $s \neq t$, $h_{i,j,s}$ is covered in **HOR**.

1013 *Proof.* At most one of the points $\{h_{i,j,t_x} : 1 \leq t_x \leq n^2\}$ and one of the points $\{v_{i,j,t_y} : 1 \leq$
 1014 $t_y \leq n^2\}$ is covered with **DIAG** (Claim 5.2).

1015 Moreover, exactly one such point h_{i,j,t_x} and one such point v_{i,j,t_y} is covered with **DIAG**,
 1016 because if none of them were covered, then the solution would have to have weight at least
 1017 $W_{\text{hv}} + 2\epsilon$ (see the proof of Lemma 5.3), which is more than $W_{\text{hv}} + k^2\delta$.

1018 We observe that points h_{i,j,t_x} and v_{i,j,t_y} have to be covered with the same segment from
 1019 **DIAG**. Indeed we need to use at least k^2 of them to use exactly one **DIAG** segment for every
 1020 pair of $1 \leq i, j \leq k$, if we used 2 segments from **DIAG** for one pair (i, j) , then we would have
 1021 used total weight at least $W_{\text{hv}} + k^2\delta + \delta$ (Lemma 5.3), which is more than $W_{\text{hv}} + k^2\delta$. Since
 1022 points h_{i,j,t_x} and v_{i,j,t_y} are covered by a single segment from **DIAG**, we have $t_x = t_y$.

1023 Therefore $t_x = t_y$ and $\text{order}(t_x)$ is an allowed tile for (i, j) because the corresponding
 1024 segment is in **DIAG**. \square

1025 We refer to the function mapping $1 \leq x \leq k$ to t_x from Lemma 5.4 as **diagonal** : $\{1, \dots, k\} \times$
 1026 $\{1, \dots, k\} \rightarrow \{1, \dots, n^2\}$.

1027 **Lemma 5.5.** *Let \mathcal{R} be any solution of a constructed instance $(\mathcal{C}, \mathcal{P}, w, 3k^2 + 2k)$ with weight*
 1028 *at most $W_{\text{hv}} + k^2\delta$. Then:*

- 1029 1. for any $1 \leq i < k, 1 \leq j \leq k$, $\text{match}_h(\text{diagonal}(i, j), \text{diagonal}(i + 1, j))$ is **true**;
- 1030 2. for any $1 \leq i \leq k, 1 \leq j < k$, $\text{match}_v(\text{diagonal}(i, j), \text{diagonal}(i, j + 1))$ is **true**.

1031 *Proof.* We prove (1) by contradiction, the proof of (2) is analogous.

1032 Let us take any $1 \leq i < k, 1 \leq j \leq k$ and name $t_1 = \text{diagonal}(i, j)$ and $t_2 = \text{diagonal}(i +$
 1033 $1, j)$. We also assume that $\text{match}_h(t_1, t_2)$ is **false**, which is equivalent to the fact that segment
 1034 $(h_{i,j,t_1}^R, h_{i+1,j,t_2}^L)$ is not in set **HOR**.

1035 Therefore h_{i,j,t_1} and h_{i+1,j,t_2} are not covered by segments from **HOR** (Lemma 5.4), while
 1036 h_{i,j,t_1}^R and h_{i+1,j,t_2}^L have to be covered by segments from **HOR** (Claim 5.1).

1037 Every segment from **HOR** either:

- 1038 • starts at point h_{x,y,z_1}^R and ends at point h_{x+1,y,z_2}^L for some $1 \leq x < k, 1 \leq y \leq k$ and
 1039 $1 \leq z_1, z_2 \leq n^2$; or
- 1040 • is $\text{horBeg}_{y,z}$ and starts at $h_{1,y,1}^L$ and ends at h_{1,y,n^2}^L for some $1 \leq y \leq k$ and $1 \leq z \leq n^2$;
 1041 or
- 1042 • is $\text{horEnd}_{y,z}$ and starts at $h_{k,y,z}^R$ and ends at h_{k,y,n^2}^R for some $1 \leq y \leq k$ and $1 \leq z \leq n^2$.

1043 All of the points between h_{i,j,t_1}^R and h_{i+1,j,t_2}^L are covered by segments in HOR and there is no
 1044 segment $(h_{i,j,t_1}^R, h_{i+1,j,t_2}^L)$ in HOR. Hence, there are at least two different segments covering
 1045 them. If both of these segments are neither $\text{horBeg}_{y,z}$ nor $\text{horEnd}_{y,z}$, then one of them must
 1046 begin at h_{i,j,t_1}^R and end at h_{i+1,j,z_2}^L and there must be other one that begins at h_{i,j,z_1}^R and ends
 1047 at h_{i+1,j,t_2}^L for some $1 \leq z_1, z_2 \leq n^2$.

1048 Thus, the space between h_{i,j,z_1}^R and $h_{i,j+1,z_2}^L$ would be covered twice and is longer than ϵ .
 1049 The case when one of them is $\text{horBeg}_{y,z}$ or $\text{horEnd}_{y,z}$ is analogous. Note that they cannot be
 1050 both $\text{horBeg}_{y,z}$ or $\text{horEnd}_{y,z}$.

1051 By the proof of Lemma 5.3, the lower bound for weight of such a solution is $W_{\text{hv}} + \epsilon$ which
 1052 is more than $W_{\text{hv}} + k^2\delta$.

1053 Therefore h_{i,j,t_1}^R and h_{i+1,j,t_2}^L must be covered by one segment from HOR, namely $(h_{i,j,t_1}^R, h_{i+1,j,t_2}^L)$.
 1054 Hence $(h_{i,j,t_1}^R, h_{i+1,j,t_2}^L)$ is a segment in HOR and $\text{match}_h(t_1, t_2)$ is **true**. \square

1055 **Lemma 5.6.** *If there exists a solution to instance $(\mathcal{C}, \mathcal{P}, w, 3k^2 + 2k)$ with weight at most*
 1056 *$W_{\text{hv}} + k^2\delta$, then there exists a solution to the grid tiling instance $(f_{i,j})$.*

1057 *Proof.* Take **diagonal** function from Lemma 5.4.

1058 To define the x function for every $1 \leq i \leq k$ set $x(i) := x_i$ where $(x_i, a) = \text{order}(v_{i,1})$.
 1059 Similarly, to define the y function, for every $1 \leq i \leq k$ set $y(i) := y_i$ where $(b, y_i) = \text{order}(h_{1,i})$

1060 To prove that this is a correct solution to grid tiling, we need to prove that for every
 1061 $1 \leq i, j \leq k$, $(x(i), y(j))$ is in the allowed tiles set $f(i, j)$.

1062 Let us take any $1 \leq i, j \leq k$. By Lemma 5.5 and simple induction, we know that
 1063 $\text{match}_h(\text{diagonal}(1, j), \text{diagonal}(i, j))$ and $\text{match}_v(\text{diagonal}(i, 1), \text{diagonal}(i, j))$ are **true**. There-
 1064 fore $\text{order}(\text{diagonal}(i, j)) = (x(i), y(j))$. By Lemma 5.4 we know that $\text{order}(\text{diagonal}(i, j))$ is in
 1065 $f(i, j)$. Therefore $(x(i), y(j))$ is in $f(i, j)$. \square

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