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Approximation and Parametrized Algorithms for Segment Set Cover

6

Master's thesis

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8

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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 The work presents a study of different geometric set cover problems. It mostly focuses on
24 segment set cover and its connection to the polygon set cover.

25

Keywords

26 set cover, geometric set cover, FPT, $W[1]$ -completeness, APX-completeness, PCP theorem,
27 NP-completeness

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36 Algorytmy parametryzowania i trudność aproksymacji problemu pokrywania zbiorów
37 odcinkami na płaszczyźnie

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Chapter 1

Introduction

The Set Cover problem is one of the most common NP-complete problems. [tutaj referencja] We are given a family of sets and have to choose the smallest subfamily of these sets that cover all their elements. This problem naturally extends to settings where we put different weights on the sets and look for the subfamily of the minimal weight. This problem is NP-complete even without weights and if we put restrictions on what the sets can be. One of such variants is Vertex Cover problem, where sets have size 2 (they are edges in a graph).

In this work we focus on another such variant where the sets correspond to some geometric shapes and only some points of the plane have to be covered. When these shapes are rectangles with edges parallel to the axis, the problem can be proven to be W[1]-complete (solution of size k cannot be found in $n^o(k)$ time), APX-complete (for sufficiently small $\epsilon > 0$, the problem does not admit $1 + \epsilon$ -approximation scheme) [referencje].

Some of these settings are very easy. Set cover with lines parallel to one of the axis can be solved in polynomial time.

There is a notion of δ -expansions, which loosen the restrictions on geometric set cover. We allow the objects to cover the points after δ -expansion and compare the result to the original setting. This way we can produce both FPT and EPTAS for the rectangle set cover with δ -extensions [referencje].

Our contribution. In this work, we prove that unweighted geometric set cover with segments is fixed parameter tractable (FPT).

Moreover, we show that geometric set cover with segments is APX-complete for unweighted axis-parallel segments, even with $1/2$ -extensions. So the problem for very thin rectangles also can't admit PTAS. Therefore, in the efficient polynomial-time approximation scheme (EPTAS) for *fat polygons* by [Har-Peled and Lee, 2009], the assumption about polygons being fat is necessary.

Finally, we show that geometric set cover with weighted segments in 3 directions is W[1]-complete. However, geometric set cover with weighted segments is FPT if we allow δ -extension.

This result is especially interesting, since it's counter-intuitive that the unweighted setting is FPT and the weighted setting is W[1]-complete. Most of such problems (like vertex cover or [wiecej przykladow]) are equally hard in both weighted and unweighted settings.

Chapter 2

Definitions

2.1. Geometric Set Cover

Define geometric set cover problem where you are given \mathcal{P} – set of objects, \mathcal{C} – set of points and you need to choose $\mathcal{R} \subset \mathcal{P}$ such that every point in \mathcal{C} is inside some element from \mathcal{R} and $|\mathcal{R}|$ is minimal.

In parametrized setting for given k , we only look among such \mathcal{R} , that $|\mathcal{R}| \leq k$.

In weighted setting, there is some given function $f : \mathcal{P} \rightarrow \mathbb{R}^+$, and we minimize $\sum_{R \in \mathcal{R}} f(R)$.

2.2. Approximation

Let's recall some of the definitions related to approximation problems that will be used in the following sections.

Definition 2.2.1 A *polynomial-time approximation scheme (PTAS)* is a family of algorithms A_ϵ for every $\epsilon > 0$, that takes an instance I of a minimization problem and in polynomial time, finds a solution that is within a factor $(1 + \epsilon)$ of being optimal. That means it has weight at most $(1 + \epsilon)\text{opt}(I)$, where $\text{opt}(I)$ is a weight of the optimal solution for I .

Definition 2.2.2 Problem is *APX-hard* if there exists such $\epsilon > 0$ so that $(1 + \epsilon)$ -approximation problem is NP-hard.

2.3. δ -extensions

Definition 2.3.1 *δ -extensions for segments* For any $\delta > 0$ and segment $L = \{(x_a, y_a), (x_b, y_b)\}$ of length $d = \sqrt{(x_a - x_b)^2 + (y_a - y_b)^2}$, a δ -extension of this segment, $L^{+\delta}$, is an open segment parallel to L , but with both ends extended by δd , i.e. if we denote vector $t = (x_b - x_a, y_b - y_a)$, $L^{+\delta} = \{(x_a, y_a) - \delta \cdot t, (x_b, y_b) + \delta \cdot t\}$

A relaxed cover problem with δ -extensions requires to cover all the points in \mathcal{C} with objects with δ -extensions (which always include no less points than the objects before δ -extensions). We also look for solution not larger than optimal solution for the original problem, so it doesn't have to be optimal for cover problem with objects after δ -extensions.

Definition 2.3.2 *Cover problem with δ -extensions* Let \mathcal{P} be an optimization cover problem that takes as an input instance I and outputs a minimal solution \mathcal{R}^{opt} , such that \mathcal{R}^{opt} covers I .

We define an optimization cover problem \mathcal{P} with δ -extensions as a problem which takes an input instance I , and outputs a solution \mathcal{R} , such that δ -extended set $\{R^{+\delta} : R \in \mathcal{R}\}$ covers I and is no worse than the optimal solution for the problem without extensions, i.e. $|\mathcal{R}| \leq |\mathcal{R}^{opt}|$.

Definition 2.3.3 PTAS with δ -extensions Let \mathcal{P} be an optimization cover problem that takes as an input instance I and outputs a minimal solution \mathcal{R}^{opt} , such that \mathcal{R}^{opt} covers I .

We define a PTAS for cover problem \mathcal{P} with δ -extensions as a problem which takes as an input instance I , and outputs a solution \mathcal{R} , such that δ -extended set $\{R^{+\delta} : R \in \mathcal{R}\}$ covers I and is within a $(1 + \epsilon)$ factor of the optimal solution for this problem without extensions, i.e. $(1 + \epsilon)|\mathcal{R}| \leq |\mathcal{R}^{opt}|$.

143 Chapter 3

144 Geometric Set Cover with segments

145 3.1. FPT for segments

146 3.1.1. Segments parallel to one of the axis

147 You can find this in Platypus book.

148 We'll show $\mathcal{O}(2^k)$ branching algorithm. Let's take point K that hasn't been covered yet
149 with the smallest coordinate in lexicographical order. We need to cover K with some of the
150 remaining segments.

151 We choose one of the 2 directions on which we will cover this point. In this direction
152 we take greedily the segment that will cover the most points (there are points in \mathcal{C} only on
153 one side of K in this direction, so all segments covering K in this direction create monotone
154 sequence of sets – zbiory zstępujące).

155 3.1.2. Segments in d directions

156 The same algorithm as before but in complexity $\mathcal{O}(d^k)$.

157 3.1.3. Segments in arbitrary direction

158 **Theorem 3.1.1 (FPT for segment cover).** *There exists an algorithm that given a family*
159 *\mathcal{P} of n segments (in any direction), a set of m points \mathcal{C} and a parameter k , runs in time*
160 *$f(k) \cdot (nm)^c$ for some computable function f and constant c , and outputs a subfamily $\mathcal{R} \subseteq \mathcal{P}$*
161 *such that $|\mathcal{R}| \leq k$ and \mathcal{R} covers all points in \mathcal{C} .*

162 **Proof.** We will show such algorithm in FPT.

163 If there exist two segments a and b in \mathcal{P} , such that any point covered by a is also covered
164 by b , then without loss of generality we can remove segment a from \mathcal{P} . We repeat this process
165 until no such (a, b) pair exists.

166 Let us first assume that we reduced our instance to a kernel, where *any line* contains no
167 more than k points.

168 Since any segment covers a set of colinear points, for such a kernel k segments can cover
169 only at most k^2 points. Therefore, for the answer to be positive, the number of points has
170 to be at most k^2 . The number of segments is now bounded by k^4 , since if we consider two
171 *extreme* points covered by a given segment, then these pairs must be distinct, otherwise two
172 segments would contain the same set of points. Since both the number of points and the

number of segments is bounded by a function of k , this instance can be easily solved in time $O(f(k))$.

It remains to show how to construct the kernel.

Assume there exists a line l containing points x_1, \dots, x_t , where $t \geq k + 1$. Note that a segment that does not lie on l can cover only at most one of the points x_i . Therefore, out of points x_1, \dots, x_{k+1} , at least one has to be covered by a segment that lies on l , let us fix x_i to be the first such point. Then, we can greedily choose a segment that lies on l , covers x_i , and also covers the largest number of points x_j for $j > i$.

Since we have at most $k + 1$ choices to branch over and each choice adds a segment to the constructed solution, we obtain an algorithm with complexity $O(k^k)$.

3.2. APX-completeness for segments parallel to axis

In this section we analyze if there exists an $(1 + \epsilon)$ -approximation scheme for set cover with rectangles. We will show that we can restrict this problem to a very easy setting: segments parallel to axes and allow $(1/2)$ -extension, and the problem is still APX-hard. Note that segments are just degenerated rectangles with one side being very narrow.

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

Theorem 3.2.1 (*axis-parallel segment set cover with $1/2$ -extension is APX-hard*). *Unweighted geometric set cover with axis-parallel segments in 2D (even with $1/2$ -extension) is APX-hard, i.e. assuming $P \neq NP$, there does not exist a PTAS for this problem.*

Theorem 3.2.1 implies the following.

Corollary 3.2.1 (*rectangle set cover is APX-hard*). *Unweighted geometric set cover with rectangles (even with $1/2$ -extension) is APX-hard.*

We will prove Theorem 3.2.1 by taking a problem that is APX-hard and showing a reduction. For this problem we choose MAX-(3,3)-SAT which we define in detail below.

Given an instance I of MAX-(3,3)-SAT, we will construct an instance J of axis-parallel segment set cover problem, such that for an sufficiently small $\epsilon > 0$, $(1 + \epsilon)$ -approximation for J will yield distinguishing whether an instance I of MAX-(3,3)-SAT is fully satisfiable or $(7/8 + \epsilon)$ -satisfiable (**Theorem 3.2.2**), which is nonapproximable beyond the random assignment assuming $P \neq NP$.

3.2.1. MAX-(3,3)-SAT problem

Definition 3.2.1 *MAX-3SAT is an optimization problem. We are given a 3-CNF formula, and need to find an assignment of variables that satisfies the most clauses.*

Definition 3.2.2 *MAX-(3,3)-SAT is a MAX-3SAT with an additional restriction that every variable appears in exactly 3 clauses, so that the number of clauses is equal to number of variables.*

In the lemmas above we use a property of MAX-(3,3)-SAT proved in [Håstad, 2001] and described in Theorem 3.2.2.

Theorem 3.2.2 [*Håstad, 2001*]

For any $\epsilon > 0$, it is NP-hard to distinguish satisfiable (3,3)-SAT formulas from $(7/8 + \epsilon)$ -satisfiable (3,3)-SAT formulas. Said equivalently, MAX-(3,3)-SAT is nonapproximable beyond the random assignment threshold on satisfiable instances.

214 The following lemma encapsulates the properties of the reduction described in this section,
 215 and it allows to prove Theorem 3.2.1.

216 **Lemma 3.2.1** *Given an instance S of MAX-(3,3)-SAT with n variables and optimal re-*
 217 *sult $OPT(S)$, we can construct an instance I of axis-parallel segments in 2D with 1/2-*
 218 *extensions, such that:*

- 219 1. *for every solution X of problem I , there exists a solution of S of size at least $15n - |X|$;*
- 220 2. *for every solution X of problem S , there exists a solution of I of size $15n - |X|$;*
- 221 3. *every solution with 1/2-extensions for I is also an solution to the original problem I .*

222 Therefore optimal solution of I is $OPT(I) = 15n - OPT(S)$.

223 We prove Lemma 3.2.1 in subsequent sections, but meanwhile let's prove Theorem 3.2.1
 224 using Lemma 3.2.1 and Theorem 3.2.2.

225 **Proof of Theorem 3.2.1** Take any $0 < \epsilon < 1/(15 \cdot 8)$.

226 Let's assume that there exists an $(1 + \epsilon)$ -approximation scheme for unweighted geometric
 227 set cover with axis-parallel segments in 2D with (1/2)-extensions. We will construct an
 228 algorithm distinguishing instances of MAX-(3,3)-SAT in Theorem 3.2.2.

229 Take an instance S of MAX-(3,3)-SAT to be distinguished and construct an instance of
 230 geometric set cover I using Lemma 3.2.1.

231 We now use the $(1 + \epsilon)$ -approximation scheme for instances of geometric set cover on I ,
 232 denote the cost of the result of this approximation as $approx(I)$.

233 We will prove that if $approx(I) \geq 15n - (\frac{7}{8} + \epsilon)n$ then S satisfied at most $(\frac{7}{8} + \epsilon)n$ clauses,
 234 and if $approx(I) < 15n - (\frac{7}{8} + \epsilon)n$ then S was satisfiable.

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

$$OPT(S) = n$$

235 From Lemma 3.2.1 we have:

$$OPT(I) = 14n$$

$$\begin{aligned} approx(I) &\leq (1 + \epsilon)OPT(I) = 14n(1 + \epsilon) = 14n + 14\epsilon \cdot n = \\ &= 14n + (15\epsilon - \epsilon)n < 14n + \left(\frac{1}{8} - \epsilon\right)n = 15n - \left(\frac{7}{8} + \epsilon\right)n \end{aligned}$$

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

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

From Lemma 3.2.1 we have:

$$OPT(I) = 15n - \left(\frac{7}{8} + \epsilon\right)n$$

236 Since approximation of problem with extensions is also approximation without extensions
 237 from Lemmma 3.2.1 3., we have:

$$approx(I) \geq OPT(I) = 15n - \left(\frac{7}{8} + \epsilon\right)n$$

238 Therefore, by using the assumed to exist $(1 + \epsilon)$ -approximation scheme for segment set
 239 cover, it is possible to distinguish S from being satisfiable and at most $(\frac{7}{8} + \epsilon)n$ satisfiable, since
 240 there exists a threshold on the approximation result in segment set cover that distinguishes
 241 these two cases. This is a contradiction, hence the approximation scheme cannot exist.

242 3.2.2. Reduction construction

243 We will show reduction from MAX-(3,3)-SAT problem to geometric set cover with segments
 244 parallel to axis. Moreover the instance of geometric set cover will be robust to 1/2-extensions
 245 (have the same optimal solution after 1/2-extension).

246 The construction will be composed of 2 types of gadgets: **VARIABLE-gadgets** and
 247 **CLAUSE-gadgets**. **CLAUSE-gadgets** would be constructed using two **OR-gadgets** con-
 248 nected together.

249 3.2.2.1. VARIABLE-gadget

250 VARIABLE-gadget is responsible for choosing a value of variable in CNF formula. It allows
 251 two minimal solution and every minimal solution must use exactly one of the (c_i, g_i) and
 252 (f_i, h_i) segments, so you can assign a binary value to the variable.

Points. Define points:

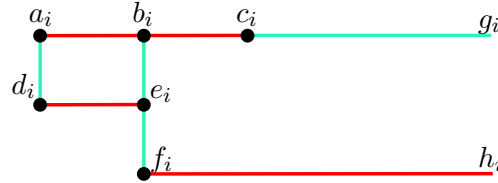


Figure 3.1: **Choose variable value gadget.** We denote set of points marked with black circle as $C_variable_i$ and need to be covered (are part of set \mathcal{C}). Note that some of the points are not marked as black dots and exists only to name segments for further reference. We denote set of red segments as x_i^{false} and set of blue segments as x_i^{true} .

253
 254 TODO: inline $L = 12n$ after finishing these formulas

$$\begin{array}{llll} a_i = (-L, 4i) & b_i = (-\frac{2}{3}L, 4i) & c_i = (-\frac{1}{3}L, 4i) & d_i = (-L, 4i + 1) \\ e_i = (-\frac{2}{3}L, 4i + 1) & f_i = (-\frac{2}{3}L, 4i + 2) & g_i = (L, 4i) & h_i = (L, 4i + 2) \end{array}$$

Let's define

$$C_variable_i = \{a_i, b_i, c_i, d_i, e_i, f_i\}$$

256 **Segments.** Let's define

$$x_i^{true} = \{(a_i, d_i), (b_i, f_i), (c_i, g_i)\}$$

$$x_i^{false} = \{(a_i, c_i), (d_i, e_i), (f_i, h_i)\}$$

$$P_variable_i = x_i^{true} \cup x_i^{false}$$

257 **Lemma 3.2.2** For any $1 \leq i \leq n$, points $C_variable_i$ can be covered using 3 segments from
258 $P_variable_i$.

259 **Proof.** We can use either set x_i^{true} or x_i^{false} .

260 **Lemma 3.2.3** For any $1 \leq i \leq n$, points $C_variable_i$ can not be covered with less than 3
261 segments from $P_variable_i$.

262 **Proof.** There is independent set $\{d_i, f_i, c_i\}$ of size 3, therefore it can not be covered with
263 less than 3 sets (segments).

264 **Lemma 3.2.4** If both segments (c_i, g_i) and (f_i, h_i) are chosen, then the covering the remain-
265 ing points from $C_variable_i$ requires at least 2 different segments from $P_variable_i$.

266 **Proof.** There is an independent set $\{a_i, e_i\}$ of size 2 in $C_variable_i - \{c_i, f_i, g_i, h_i\}$, therefore
267 it can not be covered with less than 2 sets (segments).

268 3.2.2.2. OR-gadget

269 OR-gadget has 3 important segments – $x, y, result$. x and y don't count to the weight of
270 solution of OR-gadget (they are part of different gadgets). It has a minimal solution of weight
271 w and $result$ can be chosen only if x or y are also chosen for the solution. If none of them
272 are chosen, then solution choosing $result$ segment has weight at least $w + 1$. Therefore the
273 following formula holds for a solution R assuming that R uses only w from this OR-gadget:

$$(x \in R) \vee (y \in R) \iff result \in R$$

274 **Points.**

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

$$vec_{i,j} = (10i + 3 + 3j, 4n + 2j)$$

276 Define $\{l_{i,j}, m_{i,j} \dots v_{i,j}\}$ as $\{l_0, m_0 \dots v_0\}$ shifted by $vec_{i,j}$

277 Note that $v_{i,0} = l_{i,1}$ (see Figure 3.3)

$$C_or_gadget_{i,j} = \{l_{i,j}, m_{i,j}, n_{i,j}, o_{i,j}, p_{i,j}, q_{i,j}, r_{i,j}, s_{i,j}, t_{i,j}, u_{i,j}\}$$



Figure 3.2: **OR-gadget**. We denote these point as $or_gadget_{i,j}$. We denote set of red segments as $or_{i,j}^{false}$, set of blue segments as $or_{i,j}^{true}$, green and yellow segments as $or_move_variable_{i,j}$.

278 **Segments.** We define names subsets of segments, to refer to them in lemmas.

$$or_{i,j}^{false} = \{(q_{i,j}, r_{i,j}), (s_{i,j}, u_{i,j})\}$$

$$or_{i,j}^{true} = \{(m_{i,j}, s_{i,j}), (o_{i,j}, u_{i,j}), (t_{i,j}, v_{i,j})\}$$

$$or_move_variable_{i,j} = \{(l_{i,j}, n_{i,j}), (n_{i,j}, p_{i,j})\}$$

279 Segments in OR-gadget:

$$P_or_gadget_{i,j} = or_{i,j}^{false} \cup or_{i,j}^{true} \cup or_move_variable_{i,j}$$

280 **Lemma 3.2.5** For any $1 \leq i \leq n, j \in \{0, 1\}$ and $x \in \{l_{i,j}, p_{i,j}\}$ we can cover points in
 281 $C_or_gadget_{i,j} - \{x\} \cup \{v_{i,j}\}$ with 4 segments.

282 **Proof.** We can do that using one segment from $or_move_variable_{i,j}$ (chosen depending on
 283 the value of x) and all segments from $or_{i,j}^{true}$.

284 **Lemma 3.2.6** For any $1 \leq i \leq n, j \in \{0, 1\}$, we can cover points in $C_or_gadget_{i,j}$ with 4
 285 segments from $P_or_gadget_{i,j}$.

286 **Proof.** We can do that using $or_move_variable_{i,j}$ and $or_{i,j}^{false}$.

287 3.2.2.3. CLAUSE-gadget

288 CLAUSE-gadget is responsible for calculating if choice of the variable values meets the clause
 289 in formula. It has minimal solution of weight w if at least one variable in the clause has a
 290 correct value. Otherwise it has minimal solution $w + 1$. This way by the minimal solution for
 291 the whole problem, we can tell how many clauses were satisfiable.

292 The CLAUSE-gadgets consist of two OR-gadgets. We don't want the CLAUSE-gadgets
 293 to be crammed somewhere between the very long variable segments. That's why we have a
 294 simple gadget to *pass* the value of the segment, ie. segments $(x_{i,0}, x_{i,1}), (y_{i,0}, y_{i,1}), (z_{i,0}, z_{i,1})$.
 295 Two segments and one of them is chosen if x was chosen in the solution and the other one if
 296 x wasn't.

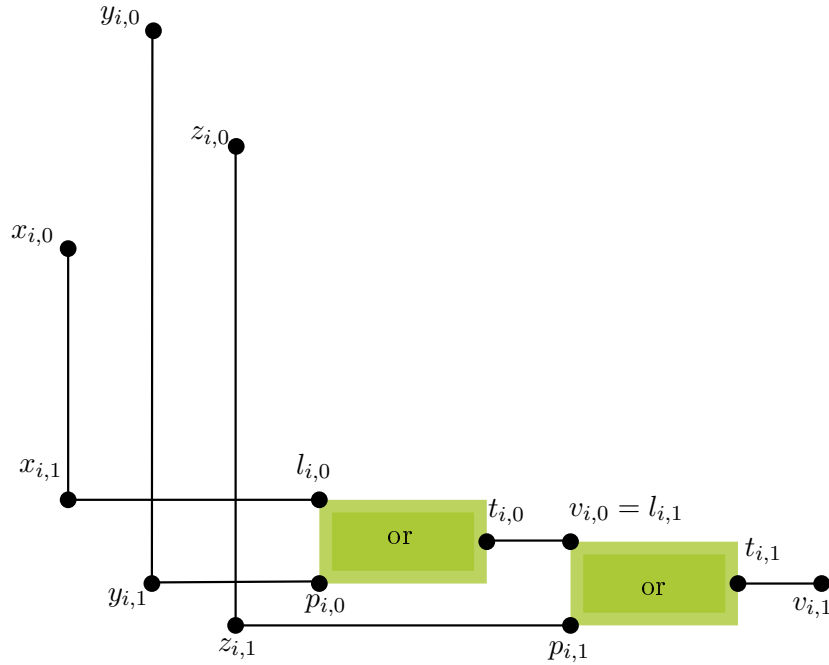


Figure 3.3: **CLAUSE-gadget**. We denote set of these points as C_clause_i . Every green rectangle is an OR-gadget. y -coordinates of $x_{i,0}$, $y_{i,0}$ and $z_{i,0}$ depend on the values of variables in the i -th clause.

297 **Points.** TODO: Rephrase it

298 Assuming clause $C_i = x_i \vee y_i \vee z_i$, function $idx(w)$ is returning index of the variable w ,
 299 function $neg(w)$ is returning whether variable w is negated in a clause.

$$\begin{aligned}
 x_{i,0} &= (10i + 1, 4 \cdot idx(x_i) + 2 \cdot neg(x_i)) & x_{i,1} &= (10i + 1, 4n) \\
 y_{i,0} &= (10i + 2, 4 \cdot idx(y_i) + 2 \cdot neg(y_i)) & y_{i,1} &= (10i + 2, 4n + 4) \\
 z_{i,0} &= (10i + 3, 4 \cdot idx(z_i) + 2 \cdot neg(z_i)) & z_{i,1} &= (10i + 3, 4n + 6)
 \end{aligned}$$

$$move_variable_i = \{x_{i,j} : j \in \{0, 1\}\} \cup \{y_{i,j} : j \in \{0, 1\}\} \cup \{z_{i,j} : j \in \{0, 1\}\}$$

$$C_clause_i = move_variable_i \cup C_or_gadget_{i,0} \cup C_or_gadget_{i,1} \cup \{v_{i,1}\}$$

Segments.

$$P_clause_i = \{(x_{i,0}, x_{i,1}), (y_{i,0}, y_{i,1}), (z_{i,0}, z_{i,1}), (x_{i,1}, l_{i,0}), (y_{i,1}, p_{i,0}), (z_{i,1}, p_{i,1}), \} \cup \\ \cup P_or_gadget_{i,0} \cup P_or_gadget_{i,1}$$

Lemma 3.2.7 For any $1 \leq i \leq n$ and $a \in \{x_{i,0}, y_{i,0}, z_{i,0}\}$, points $C_clause_i - \{a\}$ can be covered using 11 segments from P_clause_i .

Proof. For $a = x_{i,0}$ (analogous proof for $y_{i,0}$): First we use Lemma 3.2.5 twice with excluded $x = l_{i,0}$ and $x = l_{i,1} = v_{i,0}$, resulting with 8 segments $or_{i,0}^{true} \cup or_{i,1}^{true}$ 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 $a = z_{i,0}$: Using Lemma 3.2.6 and Lemma 3.2.5 with $x = p_{i,1}$, resulting with 8 segments $or_{i,0}^{false} \cup or_{i,1}^{true}$ 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: $\{(x_{i,0}, x_{i,1}), (y_{i,0}, y_{i,1}), (z_{i,1}, p_{i,1})\}$.

Lemma 3.2.8 Points C_clause_i can be covered with 12 segments from P_clause_i .

Proof. Using Lemma 3.2.6 twice we can cover $or_gadget_{i,0}$ and $or_gadget_{i,1}$ 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 3.2.9 For any $1 \leq i \leq n$, points $C_clause_i - \{x_{i,0}, y_{i,0}, z_{i,0}\}$ can not be covered using less than 11 segments from P_clause_i .

All points C_clause_i can not be covered with less than 12 segments from P_clause_i .

Proof of no cover with less than 12 segments. There is independent set of 12 points in $C_clause_i \supseteq \{x_{i,0}, y_{i,0}, z_{i,0}, l_{i,0}, p_{i,0}, q_{i,0}, u_{i,0}, v_{i,0} = l_{i,1}, p_{i,1}, q_{i,1}, u_{i,1}, v_{i,1}\}$.

Proof of no cover with less than 11 segments. We can choose disjoint sets X, Y, Z such that $X \cup Y \cup Z \subseteq C_clause_i - \{x_{i,0}, y_{i,0}, z_{i,0}\}$ and there are no segments covering points from different sets. And we will prove lower bounds for each of these sets.

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

Set X is an indendent set, so it must be covered with 3 segments.

$$Y = or_gadget_{i,0} - \{l_{i,0}, p_{i,0}\}$$

$$Z = or_gadget_{i,1} - \{l_{i,1}, p_{i,1}\}$$

For both Y and Z we can check all of the subsets of 3 segments with brutforce that none of them cover, so they have to be covered with 4 segments.

TODO: Funny fact, neither Y nor Z doesn't have independent set of size 4.

Therefore C_clause_i must be covered with at least $3 + 4 + 4 = 11$ segments.

3.2.2.4. Summary

Add some smart lemmas that sets will be exclusive to each other.

Lemma 3.2.10 Robustness to 1/2-extensions. For every segment $s \in \mathcal{P}$, s and $s^{+1/2}$ cover the same points from \mathcal{C} .

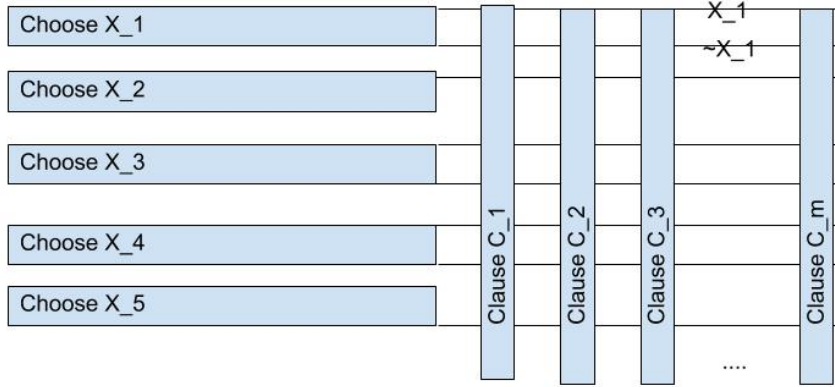


Figure 3.4: **General schema.**

General layout of VARIABLE-gadget and CLAUSE-gadget and how they interact with each other.

TODO: Rename Choose X to VARIABLE-gadget and Clause C to CLAUSE-gadget.

3.2.3. Summary of contruction

We define:

$$\mathcal{C} := \bigcup_{1 \leq i \leq n} C_variable_i \cup C_clause_i$$

$$\mathcal{P} := \bigcup_{1 \leq i \leq n} P_variable_i \cup P_clause_i$$

The subsequent sections define these sets.

We will prove some properties of different gadgets. Every segment for a gadget will only cover points in this gadget (won't interact with any diferent gadget), so we can prove lemmas *locally*.

TODO: y axis is increasing values downward on figures (not upwards like in normal).

3.2.4. Proofs of construction Lemma 3.2.1

Lemma 3.2.11 *Given an instance of MAX-(3,3)-SAT of size n with optimal solution k . For instance of geometric cover, constructed according to Lemma 3.2.1, there exists a solution of weight $15n - k$.*

Proof. Let's name the assignments of the variables in MAX-(3,3)-SAT instance, that achieve the optimal solution, $y_1, y_2 \dots y_n$, Let's cover every VARIABLE-gadget with solution described in Lemma 3.2.2, in the i -th gadget choosing the set of segments responsible for the value of y_i (true – x_i^{true} or false – x_i^{false}).

Cover every satisfied CLAUSE-gadget with solution described in Lemma 3.2.7 and unsatisfied CLAUSE-gadget with solution from Lemma 3.2.8.

This solution uses $3n + (11m + (m - k)) = 15n - k$ segments.

Lemma 3.2.12 *Given an instance of MAX-(3,3)-SAT of size n , and solution of size w to the instance of geometric cover, constructed according to Lemma 3.2.1, there exists a solution to MAX-(3,3)-SAT of size at least $15n - w$.*

Proof. Among $x_i^{true} \cup x_i^{false}$, we need to use at least 3 segments (Lemma 3.2.3). If we have chosen both segments (c_i, g_i) and (f_i, h_i) , then we have used at least 4 segments (Lemma 3.2.4).

If we chose at most one of the segments (c_i, g_i) and (f_i, h_i) , choose the corresponding variable value to the solution. If we chose both segments, choose the value that appears in most (at least 2) clauses. If we have chosen none of the segments, choose any value.

To cover $\bigcup_{1 \leq i \leq n} C_variable_i$ we have used at least $3n + a$ segments, where a is the number of i such that we have chosen both values (c_i, g_i) and (f_i, h_i) .

Among the segments responsible for the clause $C_i = x \vee y \vee z$ we need to use at least 11 segments (Lemma 3.2.9) and if we can cover it with 11 segments, then we have earlier chosen segment responsible for the value of variable x, y or z that satisfies C_i .

So we have at least 11 segments for satisfied clauses and at least 12 segments for unsatisfied clauses, so we cover it with at least $11n + b$ segments, where b is number of clauses where none of the variables x, y, z were chosen. If the segment responsible for value of x was taken, but this variable is set to have different value, then we have chosen segments for both x and $\neg x$ for this variable, so "we cheated" and this maybe clause is not met, but we assigned the value for this x_i that meets the most clauses, so for each of such "cheated" variables, at most one of the clauses isn't met.

So there are at most $a + b$ unsatisfied clauses in this instance, so we have shown the assignment with at least $n - (a + b)$ satisfied clauses.

$$w \geq 3n + a + 11n + b = 14n + a + b$$

$$15n - w \leq 15n - 14n - a - b = n - (a + b)$$

3.2.4.1. Proof of Lemma 3.2.1

Given an instance of MAX-(3,3)-SAT of size n with optimal result k . Let's construct an instance of geometric cover, constructed in aforementioned manner.

Given the Lemma 3.2.11, we know the optimal solution for the constructed geometric cover is at most $15n - k$ and since the k is optimal solution for MAX-(3,3)-SAT, then according to Lemma 3.2.12 there doesn't exist a solution with cost less than $15n - k$.

3.3. Weighted segments

3.3.1. FPT for weighted segments with δ -extensions

Theorem 3.3.1 *(FPT for weighted segment cover with δ -extensions). 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 a parameter k , runs in time $f(k) \cdot (nm)^c$ for some computable function f and constant c , and outputs a subfamily $\mathcal{R} \subseteq \mathcal{P}$ such that $|\mathcal{R}| \leq k$ and $\mathcal{R}^{+\delta}$ covers all points in \mathcal{C} .

To solve this problem we will introduce kernel for slightly different problem: Weighted segment cover of points and segments. In shortcut: WSCPS.

Lemma 3.3.1 (Algorithm for kernel of WSCPS). *There exists an algorithm that given a family \mathcal{P} of n weighted segments (in any direction), a set of m_1 points \mathcal{C}_1 and m_2 segments \mathcal{C}_2 and a parameter k , runs in time $f(k) \cdot g(m_1, m_2) \cdot n^c$ for some computable functions f, g and constant c , and outputs a subfamily $\text{sol} \subseteq \mathcal{P}$ such that $|\text{sol}| \leq k$ and sol covers all points in \mathcal{C}_1 and all segments in \mathcal{C}_2 .*

Proof Only sketch for now.

We can compute dynamic programming $dp(A, B, z)$ – the best cost to cover at least whole segment A, B using at most z segments. A, B are all interesting points – ends of any segment given on the input or points given on the input. We can compute it in polynomial time.

Then we can create a new double weighted set (original weight, number of used segments from \mathcal{P}) – \mathcal{P}_2 that has only segments which never cover partially any segment from \mathcal{C}_2 (covers the whole segment or doesn't cover at all). In such \mathcal{P}_2 we can find solution \mathcal{R} where any 2 segments have empty intersection (don't cover each other and don't meet at the ends). Because if we had such solution, we can merge these two segments and such segment there's also in \mathcal{P}_2 .

In that case we can find kernel of \mathcal{P}_2 of size $k \cdot (m_1 + 2m_2)^2$, because we only need to take the best weight covering some subset of $\mathcal{C}_1 \cup \mathcal{C}_2$.

Lemma 3.3.2 Kernel in WSCPS. *TODO: formulate it properly*

For segment cover, there is a kernel of size $f(k)$ in WSCPS.

Claim 3.3.1 *If there are more than k lines with at least $k+1$ points on them, then they can't be covered with k segments.*

Claim 3.3.2 *If there is more than k^2 points that don't lie on any line with more than k points on it, then they can't be covered with k segments.*

Claim 3.3.3 *For every long line L (with more than k points on them) we can choose $f(k)$ points on them, that if we cover all of these points with at most k segments, then the rest of the points with δ -extensions will be covered by segments in the direction of line L .*

Proof of Lemma 3.3.2. After applying the previous lemmas, we have at most $k^2 + k \cdot f(k)$ points that can be covered in any direction and for the rest of the points we can draw at most $k \cdot f(k)$ segments along their respective long lines that have to be covered by segments after δ -extensions.

Then we extend every available segment by δ -extension and we achieve the kernel in WSCPS for this instance of problem.

Lemma 3.3.3 *If all the points are covered with k segments and the biggest $2(1 + 1/\delta)^{k+1}$ spaces between points are filled, the whole segment is filled after δ -extensions of these segments.*

Proof. Let's name the $2(1+1/\delta)^{k+1}$ -st biggest space between points as y . We have guarantee that all segments of length $x > y$ are covered without δ -extensions.

Let's take one space between points that is not covered before δ -extension and we will prove it will be covered after δ -extensions. Let's assume it isn't.

This space has length x . Since it's uncovered, $x \leq y$.

Let's take side where the sum of lengths of segments covering the points is greater (left or right). Without loss of generality, let us assume it's right.

There are at most k segments to the right of this space between points. Name their lengths $l_1, l_2 \dots l_k$. If the point is covered in the other direction, the segment is degenerated to the point and $l_i = 0$. Name the space between endpoints of l_i and $l_{i+1} - x_i$. Of course, x_i is uncovered space between two points, therefore $x_i \leq y$.

TUTAJ BEDZIE PEWNIE RYSUNEK Z TYMI SUPER RZECZAMI DO PRZERW

Let's write equations meaning that i -th segment doesn't cover space x after δ -expansion.

$$l_1\delta < x \leq y \Rightarrow l_1 < y/\delta$$

$$l_2\delta < x + l_1 + x_1 < 2y + y/\delta \Rightarrow l_2 < 2y/\delta + y/\delta^2$$

$$l_3\delta < x + l_1 + x_1 + l_2 + x_2 < 3y + 3y/\delta + y/\delta^2 \Rightarrow l_3 < 3y/\delta + 3y/\delta^2 + y/\delta^3$$

From this we can "guess" induction $l_i < y((1 + 1/\delta)^i - 1)$

Trivially for $l_1 < y/\delta$.

Assume that for all $j < i$:

$$l_j < y((1 + 1/\delta)^j - 1)$$

$$\begin{aligned} l_i\delta &< x + \sum_{j=1}^{i-1} (l_j + x_j) < iy \sum_{j=1}^{i-1} l_j < iy + \sum j = 1^{i-1} y((1 + 1/\delta)^j - 1) = iy - (i - 1)y + \sum j = 1^{i-1} y(1 + 1/\delta)^j = y(1 + \sum_{j=1}^{i-1} (1 + 1/\delta)^j) = y(2 + \sum_{j=1}^{i-1} (1 + 1/\delta)^j - 1) = \\ &y(\sum_{j=0}^{i-1} (1 + 1/\delta)^j - 1) = y((1 + 1/\delta)^i / (1 - (1 + 1/\delta)) - 1) = y((1 + 1/\delta)^i \delta - 1) < y((1 + 1/\delta)^i \delta - \delta) \end{aligned}$$

Of course we also know that (since we have chosen the side with greater sum of the width of segments):

$$\sum_{i=1}^k l_i \geq 1/2 \cdot y \cdot 2(1 + 1/\delta)^{k+1} = y \cdot (1 + 1/\delta)^{k+1}$$

$$\text{But } \sum_{i=1}^k l_i < \sum_{i=1}^k y((1 + 1/\delta)^i - 1) = y((1 + 1/\delta)^{k+1} / (1 - (1 + 1/\delta)) - k) = y((1 + 1/\delta)^{k+1} \delta - k) < y(1 + 1/\delta)^{k+1}$$

Therefore the space must have been covered after δ -expansions.

3.3.2. W[1]-completeness for weighted segments in 3 directions

Theorem 3.3.2 *W[1]-completeness for weighted segments in 3 directions.* Consider the problem of covering a set \mathcal{C} of points by selecting k axis-parallel or right-diagonal weighted segments with weights from a set \mathcal{P} with minimal weight. 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 .

We will show reduction from grid tiling problem.

Let's have an instance of grid tiling problem – size of the grid k , number of elements available n and k^2 sets of available pairs in every tile $S_{i,j} \subseteq \{1, n\} \times \{1, n\}$.

448 **Construction.** We construct a set \mathcal{P} of segments and a set \mathcal{C} of points.

449 First let's choose any ordering of n^2 elements $\{1, n\} \times \{1, n\}$ and name this sequence

450 $a_1 \dots a_{n^2}$.

$$match_v(i, j) \iff a_i = \{x_i, y_i\} \wedge a_j = \{x_j, y_j\} \wedge x_i = x_j$$

$$match_h(i, j) \iff a_i = \{x_i, y_i\} \wedge a_j = \{x_j, y_j\} \wedge y_i = y_j$$

Points. Define points:

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

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

Let's define sets H and V as:

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

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

451 Let's define $\epsilon = 0.1$. For a point $\{x, y\} = p$ we define points $p^L = \{x - \epsilon, y\}$, $p^R = \{x + \epsilon, y\}$,
452 $p^U = \{x, y - \epsilon\}$, and $p^D = \{x, y + \epsilon\}$.

Then we define:

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

453 **Segments.** Define horizontal segments.

$$hor_{i,j,t_1,t_2} = (h_{i,j,t_1}^R, h_{i,j+1,t_2}^L)$$

$$ver_{i,j,t_1,t_2} = (v_{i,j,t_1}^D, v_{i,j+1,t_2}^U)$$

$$horbeg_{i,t} = (h_{i,1,1}^L, h_{i,1,t}^L)$$

$$horend_{i,t} = (h_{i,n,t}^R, h_{i,n,n^2}^R)$$

$$verbeg_{i,t} = (v_{i,1,1}^U, v_{i,1,t}^U)$$

$$verend_{i,t} = (v_{i,n,t}^D, v_{i,n,n^2}^D)$$

$$\begin{aligned} HOR &= \{hor_{i,j,t_1,t_2} : 1 \leq i \leq k, 1 \leq j < k, 1 \leq t_1, t_2 \leq n^2, match_h(t_1, t_2)\} \\ &\cup \{horbeg_{i,t} : 1 \leq i \leq k, 1 \leq t \leq n^2\} \\ &\cup \{horend_{i,t} : 1 \leq i \leq k, 1 \leq t \leq n^2\} \end{aligned}$$

$$\begin{aligned} VER &= \{ver_{i,j,t_1,t_2} : 1 \leq i \leq k, 1 \leq j < k, 1 \leq t_1, t_2 \leq n^2, match_v(t_1, t_2)\} \\ &\cup \{verbeg_{i,t} : 1 \leq i \leq k, 1 \leq t \leq n^2\} \\ &\cup \{verend_{i,t} : 1 \leq i \leq k, 1 \leq t \leq n^2\} \end{aligned}$$

$$DIAG := \{(h_{i,j,t}, v_{j,i,t}) : 1 \leq i, j \leq k, 1 \leq t \leq n^2, a_t \in S_{i,j}\}$$

TODO: explain that these segments are in fact diagonal

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

Lemma 3.3.4 *If there exists solution for grid tiling, then there exists solution for our construction using $2(k+1)k + k^2$ segments with weight exactly $2k \cdot (k(n^2 + 1) - 2 - 2\epsilon(k-1))$.*

Claim 3.3.4 *If there exists a solution to the grid tiling $c_1 \dots c_k$ and $r_1 \dots r_k$, then there exists a solution covering all points*

$$\{h_{i,j,t} : 1 \leq i, j \leq k, t = (c_i, r_j)\} \cup \{v_{i,j,t} : 1 \leq i, j \leq k, t = (c_j, r_i)\}$$

with segments in DIAG and the rest in VER or HOR and has weight $2k \cdot (k(n^2 + 1) - 2 - 2\epsilon(k-1))$.

Proof. TODO: jakiś prosty z definicji

Lemma 3.3.5 *If there exists solution for our construction using $2(k+1)k + k^2$ segments with weight exactly $2k \cdot (k(n^2 + 1) - 2 - 2\epsilon(k-1))$, then there exists a solution for grid tiling*

Proof. This follows from Lemma 3.3.6, because we just take which points are covered with DIAG.

Claim 3.3.5 *Points p^L, p^R, p^U, p^D cannot be covered with DIAG.*

Claim 3.3.6 *Points in $H \cup \{p^L : p \in H\} \cup \{p^R : p \in H\}$ cannot be covered with VER.*

Points in $V \cup \{p^U : p \in V\} \cup \{p^D : p \in V\}$ cannot be covered with HOR.

Claim 3.3.7 *For given i, j if none of the points $h_{i,j,t}$ ($v_{i,j,t}$) for $1 \leq t \leq n^2$ are covered with DIAG, then some spaces between neighbouring points were covered twice.*

Claim 3.3.8 *For given i, j two points h_{i,j,t_1}, h_{i,j,t_2} (v_{i,j,t_1}, v_{i,j,t_2}) for $1 \leq t_1 < t_2 \leq n^2$ are covered with DIAG, then one of them had to be also covered with a segment from HOR (VER).*

Proof. Point v_{i,j,t_2}^L had to be covered with VER from Claims 3.3.5 and 3.3.6. And every segment in VER covering v_{i,j,t_2}^L , covers also v_{i,j,t_1}^L .

Lemma 3.3.6 *If there exists solution for our construction with weight at most (exactly) $2k \cdot (k(n^2 + 1) - 2 - 2\epsilon(k-1))$, then for every i, j there must be exactly one t such that $h_{i,j,t}$ ($v_{i,j,t}$) is covered with DIAG and moreover if h_{i,j,t_1} and $h_{i,j+1,t_2}$ are uncovered, then $\text{math}_h(t_1, t_2)$. Analogically for v .*

Proof. Only k^2 points can be covered only in DIAG, the rest has to be covered with $VER \cup HOR$. Therefore every result must be at least $ALL_LINES - 2k^2\epsilon$, because only $2k^2$ spaces of length ϵ can be uncovered in this axis.

Of course if h_{i,j,t_1} and $h_{i,j+1,t_2}$ are uncovered, then there must exist a segment in HOR between h_{i,j,t_1}^R and $h_{i,j+1,t_2}^L$, so $\text{math}_h(t_1, t_2)$ must be true.

3.3.3. What is missing

We don't know FPT for axis-pararell segments without δ -extensions.

485 Chapter 4

486 Geometric Set Cover with lines

487 4.1. Lines parallel to one of the axis

488 When \mathcal{R} consists only of lines parallel to one of the axis, the problem can be solved in
489 polynomial time.

490 We create bipartial graph G with node for every line on the input split into sets: H –
491 horizontal lines and V – vertical lines. If any two lines cover the same point from \mathcal{C} , then we
492 add edge between them.

493 Of course there will be no edges between nodes inside H , because all of them are pararell
494 and if they share one point, they are the same lines. Similar argument for V . So the graph is
495 bipartial.

496 Now Geometric Set Cover can be solved with Vertex Cover on graph G . Since Vertex
497 Cover (even in weighted setting) on bipartial graphs can be solved in polynomial time.

498 Short note for myself just to remember how to this in polynomial time:

499 Non-weighted setting - Konig theorem + max matching

500 Weighted setting - Min cut in graph of $\neg A$ or $\neg B$ (edges directed from V to H)

501 4.2. FPT for arbitrary lines

502 You can find this is Platypus book. We will show FPT kernel of size at most k^2 .

503 (Maybe we need to reduce lines with one point/points with one line).

504 For every line if there is more than k points on it, you have to take it. At the end, if there
505 is more than k^2 points, return NO. Otherwise there is no more than k^4 lines.

506 In weighted settings among the same lines with different weights you leave the cheapest
507 one and use the same algorithm.

508 4.3. APX-completeness for arbitrary lines

509 We will show a reduction from Vertex Cover problem. Let's take an instance of the Vertex
510 Cover problem for graph G . We will create a set of $|V(G)|$ pairwise non-pararell lines, such
511 that no three of them share a common point.

512 Then for every edge in $(v, w) \in E(G)$ we put a point on crossing of lines for vertices v
513 and w . They are not pararell, so there exists exactly one such point and any other line don't
514 cover this point (any three of them don't cross in the same point).

Solution of Geometric Set Cover for this instance would yield a sound solution of Vertex Cover for graph G . For every point (edge) we need to choose at least one of lines (vertices) v or w to cover this point.

Vertex Cover for arbitrary graph is APX-complete, so this problem is also APX-complete.

4.4. 2-approximation for arbitrary lines

Vertex Cover has an easy 2-approximation algorithm, but here very many lines can cross through the same point, so we can do d -approximation, where d is the biggest number of lines crossing through the same point. So for set where any 3 lines don't cross in the same point it yields 2-approximation.

The problematic cases are where through all points cross at least k points and all lines have at least k points on them. It can be created by casting k -grid in k -D space on 2D space.

Greedy algorithm yields $\log |\mathcal{R}|$ -approximation, but I have example for this for bipartial graph and reduction with taking all lines crossing through some point (if there are no more than k) would solve this case. So maybe it works.

Unfortunately I haven't done this :(

I can link some papers telling it's hard to do.

4.5. Connection with general set cover

Problem with finite set of lines with more dimensions is equivalent to problem in 2D, because we can project lines on the plane which is not perpendicular to any plane created by pairs of (point from \mathcal{C} , line from \mathcal{P}).

Of course every two lines have at most one common point, so is every family of sets that have at most one point in common equivalent to some geometric set cover with lines?

No, because of Desargues's theorem. Have to write down exactly what configuration is banned.

539 Chapter 5

540 Geometric Set Cover with polygons

541 5.1. State of the art

542 Covering points with weighted discs admits PTAS [Li and Jin, 2015] and with fat polygons
543 with δ -extensions with unit weights admits EPTAS [Har-Peled and Lee, 2009].

544 Although with thin objects, even if we allow δ -expansion, the Set Cover with rectangles is
545 APX-complete (for $\delta = 1/2$), it follows from APX-completeness for segments with δ -expansion
546 in Section 3.2.

547 Covering points with squares is W[1]-hard [Marx, 2005]. It can be proven that assuming
548 *SETH*, there is no $f(k) \cdot (|\mathcal{C}| + |\mathcal{P}|)^{k-\epsilon}$ time algorithm for any computable function f and
549 $\epsilon > 0$ that decides if there are k polygons in \mathcal{P} that together cover \mathcal{C} , *Theorem 1.9* in [Marx
550 and Pilipczuk, 2015].

⁵⁵¹ Chapter 6

⁵⁵² Conclusions

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