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# Parameterized Complexity of the Sparsest $k$ -Subgraph Problem in Chordal Graphs

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**Abstract.** In this paper we study the SPARSEST  $k$ -SUBGRAPH problem which consists in finding a subset of  $k$  vertices in a graph which induces the minimum number of edges. The SPARSEST  $k$ -SUBGRAPH problem is a natural generalization of the INDEPENDENT SET problem, and thus is  $\mathcal{NP}$ -hard (and even  $W[1]$ -hard) in general graphs. In this paper we investigate the parameterized complexity of both SPARSEST  $k$ -SUBGRAPH and DENSEST  $k$ -SUBGRAPH in chordal graphs. We first provide simple proofs that DENSEST  $k$ -SUBGRAPH in chordal graphs is FPT and does not admit a polynomial kernel unless  $\mathcal{NP} \subseteq \text{co}\mathcal{NP}/\text{poly}$  (both parameterized by  $k$ ). More involved proofs will ensure the same behavior for SPARSEST  $k$ -SUBGRAPH in the same graph class. We lastly provide an FPT algorithm in interval graphs for SPARSEST  $k$ -SUBGRAPH, but parameterized by the number of edges of the solution (a stronger parameterization than by  $k$ ).

## 1 Introduction

*Presentation of the problem.* Given a simple undirected graph  $G = (V, E)$  and an integer  $k$ , the SPARSEST  $k$ -SUBGRAPH problem asks to find  $k$  vertices in  $G$  inducing the minimum number of edges. The decision version asks if there exists a  $k$ -subgraph inducing at most  $C$  edges. As a generalization of the classical INDEPENDENT SET problem (for  $C = 0$ ), SPARSEST  $k$ -SUBGRAPH is  $\mathcal{NP}$ -hard in general graphs, as well as  $W[1]$ -hard when parameterized by  $k$  (as INDEPENDENT SET is  $W[1]$ -hard [10]). In addition, there is an obvious  $XP$  algorithm for SPARSEST  $k$ -SUBGRAPH when parameterized by  $k$ , as all subsets of size  $k$  can be enumerated in  $\mathcal{O}(n^k)$  time, where  $n$  is the number of vertices in the graph.

*Related problems.* Several problems closely related to SPARSEST  $k$ -SUBGRAPH have been extensively studied in the past decades. Among them, one can mention the maximization version of SPARSEST  $k$ -SUBGRAPH, namely the DENSEST  $k$ -SUBGRAPH, for which several results have been obtained in general or restricted graphs. In [8], the authors showed that DENSEST  $k$ -SUBGRAPH remains  $\mathcal{NP}$ -hard in bipartite, comparability and chordal graphs, and is polynomial-time solvable in trees, cographs, and split graphs. The complexity status of DENSEST  $k$ -SUBGRAPH in interval graphs, proper interval graphs and planar graphs is left as an open problem, and is still not answered yet. More recently, [5] improved some of these results by showing that both DENSEST  $k$ -SUBGRAPH and SPARSEST  $k$ -SUBGRAPH are polynomial-time solvable in bounded cliquewidth graphs, and [3] developed exact algorithms for SPARSEST  $k$ -SUBGRAPH, DENSEST  $k$ -SUBGRAPH and other similar problems in general graphs parameterized by  $k$  and the maximum degree  $\Delta$  of the graph. During the past two decades, a large amount of work has been dedicated to the approximability of DENSEST

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$k$ -SUBGRAPH in general graphs. So far, the best approximation ratio is  $O(n^\delta)$  for some  $\delta < 1/3$  [9], while the only negative result is due to Khot [14] ruling out a PTAS under some complexity assumptions. Still concerning DENSEST  $k$ -SUBGRAPH but in restricted graph classes, [16] developed a PTAS in interval graphs, and [7, 15] developed constant approximation algorithms in chordal graphs. In [17], we recently proved that SPARSEST  $k$ -SUBGRAPH remains  $\mathcal{NP}$ -hard in chordal graphs and admits a 2-approximation algorithm. We can also mention the dual version of SPARSEST  $k$ -SUBGRAPH, namely the MAXIMUM PARTIAL VERTEX COVER problem, for which we are looking for  $k$  vertices in the input graph which *cover* the maximum number of edges. Very recently [1] and [13] independently proved the  $\mathcal{NP}$ -hardness of MAXIMUM PARTIAL VERTEX COVER in bipartite graphs, which directly transfers to SPARSEST  $k$ -SUBGRAPH (since finding  $k$  vertices covering the maximum number of edges is equivalent to find  $(n - k)$  vertices inducing the minimum number of edges).

More generally, SPARSEST  $k$ -SUBGRAPH, DENSEST  $k$ -SUBGRAPH and MAXIMUM PARTIAL VERTEX COVER fall into the family of *cardinality constrained optimization problems* introduced by Cai [6]. In its survey, the author proved that these three problems are  $W[1]$ -hard in regular graphs, and gives an  $XP$  algorithm for general graphs with a better running time than the trivial algorithm.

As said previously, SPARSEST  $k$ -SUBGRAPH and DENSEST  $k$ -SUBGRAPH are natural generalizations of  $k$ -INDEPENDENT SET and  $k$ -CLIQUE, and are thus important both from a theoretical and practical point of view. Our motivation is to study their computational (parameterized) complexity in graph classes where they remains  $\mathcal{NP}$ -hard whereas  $k$ -INDEPENDENT SET and  $k$ -CLIQUE are polynomial-time solvable, such as the well-known class of perfect graphs and some of its subclasses. To that end, we study their parameterized complexity in the class of chordal graphs, an important subclass of perfect graphs which arises in many practical situations [12]. More precisely, we prove that both SPARSEST  $k$ -SUBGRAPH and DENSEST  $k$ -SUBGRAPH in chordal graphs are fixed-parameter tractable and do not admit a polynomial kernel under some classical complexity assumptions. As we will see, the results are quite easy to obtain for DENSEST  $k$ -SUBGRAPH, but require some efforts for SPARSEST  $k$ -SUBGRAPH.

*Organization of the paper.* The paper is organized as follows: in the following section (Section 2), we recall the classical definitions of parameterized complexity and chordal graphs. Our two main results, namely the  $FPT$  algorithm and kernel lower bound for SPARSEST  $k$ -SUBGRAPH in chordal graphs, are presented respectively in Sections 4 and 5. Lastly in Section 6, we present an  $FPT$  algorithm for SPARSEST  $k$ -SUBGRAPH in interval graphs but parameterized by the number of edges in a solution, a stronger parameter than  $k$  (since it is always smaller than  $\binom{k}{2}$ ). Before all these, we study as an appetizer the parameterized complexity of DENSEST  $k$ -SUBGRAPH in chordal graphs in Section 3.

## 2 Parameterized Algorithms, Chordal Graphs, Interval Graphs

*Parameterized algorithms.* An interesting way to tackle  $\mathcal{NP}$ -hard problems is parameterized complexity. A parameterized problem  $Q$  is a subset of  $\Sigma^* \times \mathbb{N}$ , where the second component is called the *parameter* of the instance. A *fixed-parameter tractable* ( $FPT$  for short) problem is a problem for which there exists an algorithm which, given  $(x, k) \in \Sigma^* \times \mathbb{N}$ , decides whether  $(x, k) \in Q$  in time  $f(k)|x|^{O(1)}$  for some computable function  $f$ . Such an algorithm becomes efficient with an hopefully small parameter. A *kernel* is a polynomial algorithm which, given  $(x, k) \in \Sigma^* \times \mathbb{N}$ , outputs an instance  $(x', k')$  such that  $(x, k) \in Q \Leftrightarrow (x', k') \in Q$  and  $|x'| + k' \leq f(k)$  for some computable function  $f$ . The existence of a kernel is equivalent to the existence of an  $FPT$ -algorithm. Nevertheless one can

ask the function  $f$  to be a polynomial. If so, then the kernel is called a *polynomial kernel*. If a problem admits a polynomial kernel, then it roughly means that we can, in polynomial time, compress the initial instance into an instance of size  $\text{poly}(k)$  which contains all the hardness of the instance. In order to rule out polynomial kernels, we will use the recent technique of cross-composition [2].

Roughly speaking, a cross-composition is a polynomial reduction from  $t$  instances of a (non-parameterized) problem  $A$  to a single instance of a parameterized problem  $B$  such that the constructed instance is positive iff one of the input instances is positive. In addition, the parameter of the constructed instance must be of size polynomial in the maximum size of the input instances and the logarithm of  $t$ . It is known that if  $A$  is  $\mathcal{NP}$ -hard and  $A$  cross-composes into  $B$ , then  $B$  cannot admit a polynomial kernel under some complexity assumptions. For a stronger background concerning the parameterized complexity, we refer the reader to [10]. Formal definitions of cross-composition and related notions are available in Appendix A.

*Chordal graphs.* A graph  $G = (V, E)$  is a chordal graph if it does not contain an induced cycle of length at least four. As said previously, chordal graphs form an important subclass of perfect graphs. One can also equivalently define chordal graphs in terms of a special tree decomposition. Indeed, it is known [11] that a graph  $G = (V, E)$  is a chordal graph if and only if one can find a tree  $T = (\mathcal{X}, A)$  with  $\mathcal{X} \subseteq 2^V$  such that for all  $v \in V$ , the set of nodes of  $T$  containing  $v$ , that is  $\mathcal{X}_v = \{X \in \mathcal{X} : v \in X\}$ , induces a (connected) tree, and such that for all  $u, v \in V$  we have  $\{u, v\} \in E$  if and only if  $\mathcal{X}_u \cap \mathcal{X}_v \neq \emptyset$ . Moreover, given a chordal graph, this corresponding tree can be found in polynomial time. From this definition, it is clear that each  $X \in \mathcal{X}$  induces a clique in  $G$ .

*Interval Graphs.* An interesting subclass of chordal graphs is the class of interval graphs. A graph  $G = (V, E)$  is a interval graph if there exists an interval model of  $G$ , *i.e.* if each vertex  $v \in V$  can be mapped to an interval  $I_v$  of the real line, such that two intervals overlap if and only if their corresponding edges are adjacent. It is easy to see that for each interval graphs, there exists a tree decomposition  $T = (\mathcal{X}, A)$  such that  $T$  is a path.

### 3 Appetizer: Parameterized Complexity of Densest $k$ -Subgraph in Chordal Graphs

**Theorem 1.** *DENSEST  $k$ -SUBGRAPH in chordal graphs is FPT and does not admit a polynomial kernel unless  $\mathcal{NP} \subseteq \text{coNP}/\text{poly}$ , both parameterized by  $k$ .*

Indeed, notice that if the input graph  $G$  contains a clique of size  $k$  or more (which can be tested in polynomial time in chordal graphs), then it must be an optimal solution. Otherwise, it implies that the treewidth of  $G$  is upper bounded by  $k$  (since the treewidth equals the maximum clique number in chordal graphs), and we can apply the dynamic programming of [4] over a classical tree decomposition of  $G$  in order to compute an optimal solution in *FPT* time.

For the kernel lower bound, we can easily construct a cross-composition from itself, by taking disjoint union of  $t$  graphs on  $n$  vertices  $G_1, \dots, G_t$ , and adding  $n^2$  universal vertices to each graph.

## 4 FPT Algorithm for Sparsest $k$ -Subgraph in Chordal Graphs

**Definitions and Notations.** Let  $G = (V, E)$  be a chordal graph and  $T = (\mathcal{X}, A)$  be its corresponding tree decomposition as defined in section 2. Recall that for each  $X \in \mathcal{X}$ ,  $X$  induces a clique in  $G$ .

We denote respectively by  $\mathcal{L}$  and  $\mathcal{I}$  the set of leaves and internal nodes of  $T$  (we have  $\mathcal{X} = \mathcal{L} \cup \mathcal{I}$ ). In the following we suppose that  $T$  is rooted at an arbitrary node  $X_r$ . Let  $X \in \mathcal{X}$ , we denote by  $pred(X)$  the unique predecessor of  $X$  in  $T$  (by convention  $pred(X_r) = \emptyset$ ), and by  $succ(X)$  the set of successors of  $X$  in  $T$ . For a vertex  $v \in V$  (resp. a node  $X \in \mathcal{X}$ ), we denote by  $d(v)$  (resp.  $d(X)$ ) its degree in  $G$  (resp. in  $T$ ). For a set of vertices  $U \subseteq V$  (resp. set of nodes  $A \subseteq \mathcal{X}$ ), we denote by  $G[U]$  (resp.  $T[A]$ ) the subgraph of  $G$  induced by  $U$  (resp. the subforest of  $T$  induced by  $A$ ). We say that a vertex  $v \in V$  is a *lonely*<sup>1</sup> vertex (resp. *almost lonely* vertex) if  $|\mathcal{X}_v| = 1$  (resp.  $|\mathcal{X}_v| = 2$ ), i.e. if it appears in only one (resp. two) nodes of  $T$ .

**First Observations.** A maximum independent set can be computed in polynomial time in chordal graphs (since chordal graphs are perfect). Hence, we first determine if there exists an independent set of size  $k$ . In this case, we return this set which is naturally an optimal solution. Thus, we assume in the following that the graph  $G$  does not contain an independent set of size  $k$ .

Notice that we can assume that for every leaf  $L$  of the tree we do not have  $L \subseteq pred(L)$  (otherwise we can contract the two nodes). Therefore, for each leaf  $L$  of the tree, there is a vertex  $x \in L$  such that  $x \notin pred(L)$ , i.e.  $x$  is a lonely vertex. Since there is no independent set of size  $k$  in  $G$ , and since lonely vertices of leaves are pairwise non adjacent, we have the following:

**Observation 1** *We can assume that  $|\mathcal{L}| < k$ .*

Let us now state a simple property verified by optimal solutions. Let  $S$  be a set of  $k$  vertices. Assume that there are vertices  $x \in S$  and  $y \in V \setminus S$  such that  $\mathcal{X}_y \subsetneq \mathcal{X}_x$ . Then it means that  $N(y) \subseteq N(x)$ . Thus, if we replace  $x$  by  $y$  in the solution, the number of edges in the solution cannot increase. A set  $S$  is *closed under inclusion* if there is no vertex  $x$  in  $S$  such that there exists  $y \in V \setminus S$  such that  $\mathcal{X}_y \subsetneq \mathcal{X}_x$ . So there always exists an optimal solution closed under inclusion.

**Idea of the Algorithm.** Our goal is to find an optimal solution closed under inclusion. First note that any optimal solution closed under inclusion must contain a lonely vertex per leaf of  $T$ . Indeed, as each leaf  $L$  is not included in  $pred(L)$ , there exists a lonely vertex  $x$  in  $L$ . Thus, either the solution intersects  $L$ , and since it is closed by inclusion it contains a lonely vertex, or we can take a vertex of the solution and replace it by  $x$ , which does not create any additional edge (since no vertex of  $N(x) = L \setminus \{x\}$  was in the solution).

Our method can be summarized as follows. First, we take a lonely vertex in each leaf and guess a binary flag  $w(L) \in \{0, 1\}$  for each leaf  $L$  which indicates whether another vertex of  $L$  has to (with value 1) or does not have to (with value 0) be taken in the solution. The width of such a branching is bounded according to Observation 1. Then, given a leaf  $L$  with  $w(L) = 1$ , we first try to add to the solution the “most interesting” vertex of the leaf (for example a lonely vertex). When this is

<sup>1</sup> Notice that every lonely vertex is a so-called “simplicial vertex” (a vertex whose neighborhood is a clique). However, if a node of the tree is contained in another node, a simplicial vertex may not be a lonely one. Since we do not make any supposition on the tree  $T$  (we will in particular duplicate nodes during the algorithm), we will prefer the term “lonely”.

not possible (the neighborhood of the vertices of  $L$  can be incomparable if these vertices appear on incomparable subtrees), we apply some branching rules that re-structure the tree and create new “interesting vertices”.

**Terminology for the Algorithm.** The algorithm is a branching algorithm composed of pre-processing rules (which do not require branching) and branching rules. When a rule is applied, we assume that previous rules cannot be applied.

During the algorithm, a partial solution  $S$  (initialized to  $\emptyset$ ) will be constructed, and the input graph  $G = (V, E)$  together with  $k$ ,  $T$  (and thus  $\mathcal{X}$ ,  $\mathcal{L}$  and  $\mathcal{I}$ ) will be modified. To avoid heavy notation we will keep these variables to denote the current input, and denote by  $G_0$  the original graph, and by  $N_0$  the neighborhood function of  $G_0$ .

In the following, *taking* a vertex  $v \in V$  in the solution means that  $v$  is added to  $S$ , and  $v$  is removed both from the graph  $G$  and the tree  $T$  (removing each of its occurrences). *Deleting* a vertex means removing the vertex from  $G$  and from  $T$ . If a leaf of  $T$  becomes empty after taking or deleting a vertex, then simply remove the leaf.

Let  $F \in \mathcal{I}$  be a leaf of  $T[\mathcal{I}]$  (*i.e.* a node of  $T$  of which all successors are leaves). The node  $F$  is a *bad father* if there exists a vertex  $u$  which appears in at least two leaves of  $\text{succ}(F)$ . So a node is a bad father if the leaves attached to it are not vertex disjoint. We denote by  $\#BF$  the number of bad fathers of the tree. Finally, we denote by  $\#AL$  (for “almost leaf”) the number of internal nodes of  $T$  such that at least one successor is a leaf. Notice that  $\#AL, \#BF \leq |\mathcal{L}|$ .

In addition, as said previously, we will put “flags” on some leaves  $\mathcal{L}^* \subseteq \mathcal{L}$  by introducing a boolean function  $w : \mathcal{L}^* \rightarrow \{0, 1\}$ , which indicates whether it intersects the solution (value 1) or not (value 0). At the beginning of the algorithm we have  $\mathcal{L}^* = \emptyset$ . For a solution  $S \subseteq V_0$ , we say that  $S$  *respects* the flags  $w$  if for all  $L \in \mathcal{L}^*$ ,  $w(L) = 0$  iff  $S \cap L = \emptyset$ . During the algorithm we will use the term “guessing” the value  $w(L)$  of  $L \in \mathcal{L}$ . By this, we mean that we try the two possible choices (consistent with the previous ones), creating at most two distinct executions of the algorithm. Notice that  $\mathcal{L}^*$  will be implicitly updated (*i.e.*  $L$  belongs to  $\mathcal{L}^*$  in the next executions if we have guessed  $w(L)$ ).

We also add a function  $g : \mathcal{L}^* \rightarrow 2^S$ . Roughly speaking, we will modify  $g$  during the algorithm such that  $g$  remembers the neighbors of the remaining vertices  $V$  in the partial solution  $S$  already constructed. Notice that we introduced  $g$  only for the analysis, and more precisely for maintaining our invariants (see below).

**Correctness and Time Complexity.** As usually in a branching algorithm, we bound the time complexity by bounding both the depth and the maximum degree of the search tree. More precisely, we will show that:

- Each rule can be applied in FPT time.
- The branching degree of each branching rule is a function of  $k$ .
- Any branching rule strictly decreases  $(k, \#AL, \#BF)$  using the lexicographic order, whose initial value only depends on the initial value of  $k$  (by Observation 1).
- Any pre-processing rule does not increase  $(k, \#AL, \#BF)$  and decreases  $|V| + |\mathcal{I}|$ .

Thus, the number of branching steps of the search tree is a function of  $k$  only, whereas the number of steps between two branchings is polynomial in  $n$  (recall that  $|\mathcal{X}|$  is polynomial in  $n$ ), which leads to an *FPT* running time.

Recall that  $S$  denotes the partial current solution. Concerning the correctness of the algorithm, we will say that a rule is *safe* if it preserves all the following invariants:

1. The tree  $T$  is still a tree decomposition (as defined in 2) of  $G$ , which is an induced subgraph of  $G_0$ .
2. If a vertex of the partial solution is adjacent to a "surviving" vertex  $v \in V$ , then  $v$  must appear in a leaf where a flag is defined, *i.e.*:  

$$N_0(S) \cap V \subseteq \bigcup_{L \in \mathcal{L}^*} L.$$
3. The neighborhood of a surviving vertex  $u$  in the partial solution is defined by the union of  $g(L)$  for each  $L$  in which  $u$  appears, *i.e.*  $g : \mathcal{L}^* \rightarrow 2^S$  is such that  $\forall u \in V$  we have  $N_0(u) \cap S = \bigcup_{L \in \mathcal{X}_u \cap \mathcal{L}^*} g(L)$ .  
 In particular, this invariant implies that if there are  $u, v \in V$  such that  $\mathcal{X}_u \cap \mathcal{L}^* \subseteq \mathcal{X}_v \cap \mathcal{L}^*$  (*i.e.*  $v$  appears in at least as many labelled leaves as  $u$ ), then we must have  $N_0(u) \cap S \subseteq N_0(v) \cap S$  (*i.e.*  $v$  is adjacent to at least as many vertices of the solution as  $u$ ).
4. If there is an optimal solution (closed under inclusion)  $S^* \subseteq V$  such that  $S \subseteq S^*$ , and  $S^*$  respects the flags  $w$ , then one of the branching will output an optimal solution.

**Reduction Rules.** Notice that each of the following rules defines a new value for variables  $k$ ,  $T$ ,  $S$ ,  $w$  and  $g$ . However, for the sake of readability we will not mention variables that are not modified.

***Pre-Processing Rule 1: useless duplicated node.***

If there exists  $X \in \mathcal{X}$  such that  $X \notin \mathcal{L}^*$  and  $X \subseteq \text{pred}(X)$ , then contract  $X$  and  $\text{pred}(X)$  (*i.e.* delete  $X$ , and connect every  $Y \in \text{succ}(X)$  to  $\text{pred}(X)$ ).

**Lemma 1.** *Pre-Processing Rule 1 is safe.*

*Proof.* The new tree still verifies invariant 1. As  $X \notin \mathcal{L}^*$ ,  $\bigcup_{L \in \mathcal{L}^*} L$  remains unchanged, and since  $S$  is also unchanged, invariant 2 is clearly preserved. In the same way, as  $X \notin \mathcal{L}^*$ ,  $\bigcup_{L \in \mathcal{X}_u \cap \mathcal{L}^*} g(L)$  remains unchanged for any  $u$ , and invariant 3 is preserved. Invariant 4 remains true as we do not modify  $S$  nor  $w$ .

Let us now check what is decreasing when applying this rule. Notice that this rule may increase  $\#BF$  as  $\text{pred}(X)$  may become a bad father. However, in this case the rule decreases  $\#AL$ , and thus  $(k, \#AL, \#BF)$  decreases. Otherwise (if  $\#BF$  does not increase), either  $\#AL$  decreases, or  $(k, \#AL, \#BF)$  remains unchanged and  $|V| + |\mathcal{I}|$  decreases.  $\square$

***Pre-Processing Rule 2: removing a (almost) lonely vertex.***

If there exists  $L \in \mathcal{L}^*$  such that  $w(L) = 0$ , then if  $L$  contains a lonely vertex  $v$ , delete  $v$ . Otherwise, if  $L$  contains an almost lonely vertex  $v$ , then delete  $v$ .

**Lemma 2.** *Pre-Processing Rule 2 is safe.*

*Proof.* Here again invariant 1 still holds. Then, since we just remove a vertex from the graph and do not modify the solution, invariant 2 is still true. For the same reason, and since  $g(L)$  is not modified either, invariant 3 holds. Let us prove that invariant 4 is preserved. Consider an optimal solution  $S^*$  closed under inclusion which satisfies  $S \subseteq S^*$  and the flags on the leaves. As  $w(L) = 0$ , no vertex of  $L$  is used in  $S^*$ , and in particular  $v \notin S^*$ . Thus, vertices of  $S^* \setminus S$  are still in the remaining graph and the invariant still holds.

Then, obviously  $|V| + |\mathcal{I}|$  decreases while  $k$  remains unchanged. The only case in which  $\#BF$  may increase is when  $L = \{v\}$  and  $\text{succ}(\text{pred}(L)) = L$  (*i.e.*  $L$  was the unique leaf of  $\text{pred}(L)$ ), and  $\text{pred}(\text{pred}(L))$  is an almost leaf). In this case  $L$  is deleted and thus  $\text{pred}(L)$  now becomes a leaf

and  $\text{pred}(\text{pred}(L))$  may become a bad father. However in this case  $\text{pred}(L)$  and  $\text{pred}(\text{pred}(L))$  were two almost leaves, and thus the deletion of  $v$  (and thus  $L$ ) decreases  $\#AL$ , which proves that  $(k, \#AL, \#BF)$  cannot increase.  $\square$

**Branching Rule 1: taking a lonely vertex.**

If there exists  $L \in \mathcal{L}^*$  such that  $w(L) = 1$  and  $L$  contains a lonely vertex  $v$ , then take  $v$  in the solution and decrease  $k$  by one. In addition, add  $v$  into  $g(L)$ , and if  $L$  does not become empty, then guess a new value  $w(L)$ .

**Lemma 3.** *Branching Rule 1 is safe.*

*Proof.* Here again a vertex is deleted from the graph and thus invariant 1 is still verified. In addition, neighbors of  $v$  in the remaining graph must appear in the leaf  $L$  only (since  $v$  is lonely), which receives a flag  $w(L)$ . Hence invariants 2 holds. Since  $v$  has been added to  $g(L)$ , invariant 3 holds too. For the last invariant, let us consider  $S^*$  an optimal solution closed under inclusion such that  $S \subseteq S^*$  and  $S^*$  satisfies the flags of  $w$ . Suppose that  $v \notin S^*$ . Let  $x \in L \cap S^*$  (such a vertex must exist, according to  $w(L)$ ). Let us prove that  $N_0(v) \cap S^* \subseteq N_0(x) \cap S^*$  (as this will imply that replacing  $x$  by  $v$  in  $S^*$  cannot increase its cost). By invariant 3, it holds that  $N_0(v) \cap S \subseteq N_0(x) \cap S$ . By invariant 1 and by definition of the tree  $T$ , it holds that  $N_0(v) \cap S^* \cap V \subseteq N_0(x) \cap S^* \cap V$ . Since  $S^* = S \cup (S^* \cap V)$ , the result follows and invariant 4 is true.

Finally, it is clear that  $k$  decreases.  $\square$

*Remark 1.* At this point, since *Pre-Processing Rule 1* does not apply, it is clear that every leaf  $L \in \mathcal{L} \setminus \mathcal{L}^*$  contains a lonely vertex. The following branching rule aims to process these leaves.

**Branching Rule 2: processing leaves with no flag.**

If there exists  $L \in \mathcal{L} \setminus \mathcal{L}^*$ , then take a lonely vertex  $v \in L$  in the solution and decrease  $k$  by one. In addition, add  $v$  into  $g(L)$ , and if  $L$  does not become empty, guess a value  $w(L)$ .

**Lemma 4.** *Branching Rule 2 is safe.*

*Proof.* Using the same arguments as in *Branching Rule 1*, invariants 1, 2 and 3 hold. Then, let us consider  $S^*$  an optimal solution closed under inclusion such that  $S \subseteq S^*$  and  $S^*$  satisfies the flags of  $w$ . Suppose that  $v \notin S^*$ . Since  $L$  has no flag, two cases may happen: either  $S^* \cap L = \emptyset$  or  $S^* \cap L \neq \emptyset$ . In the first case, since invariant 2 implies  $N_0(v) \cap S = \emptyset$ , and since  $v$  is a lonely vertex, we have  $N_0(v) \cap S^* = \emptyset$ . Hence replacing any other vertex of  $S^*$  by  $v$  cannot increase its number of induced edges. Suppose now that  $S^* \cap L \neq \emptyset$ , and let  $x \in L \cap S^*$ . As in the proof of *Branching Rule 1*, let us prove that  $N_0(v) \cap S^* \subseteq N_0(x) \cap S^*$  (as this will imply that replacing  $x$  by  $v$  in  $S^*$  cannot increase its cost). By invariant 3, it holds that  $N_0(v) \cap S \subseteq N_0(x) \cap S$ . By invariant 1 and by definition of the tree  $T$ , it holds that  $N_0(v) \cap S^* \cap V \subseteq N_0(x) \cap S^* \cap V$ . Since  $S^* = S \cup (S^* \cap V)$ , the result follows and invariant 4 is true.

Here again it is clear that  $k$  decreases.  $\square$

*Remark 2.* At this point, notice that  $\mathcal{L}^* = \mathcal{L}$ , i.e. a flag has been assigned to each leaf. Indeed, suppose that there exists  $L \in \mathcal{L} \setminus \mathcal{L}^*$ . If  $L$  contains a lonely vertex, then *Branching Rule 1* must apply. Otherwise, *Pre-Processing Rule 1* must apply. In addition, there is no lonely vertex in the leaves, as otherwise *Branching Rule 1* or *Pre-Processing Rule 2* would apply.



**Branching Rule 3: partitioning leaves of a bad father.**

If there exists a bad father  $F \in \mathcal{X}$ , let  $\mathcal{L}'$  be the set of leaves in  $\text{succ}(F)$  and  $C = \bigcup_{L \in \mathcal{L}'} L$  be the set of vertices contained these leaves. Partition  $C$  into the equivalence classes  $C_1, \dots, C_t$  of the following equivalence relation: two vertices  $u, v \in C$  are equivalent if  $\mathcal{X}_u \cap \mathcal{L}' = \mathcal{X}_v \cap \mathcal{L}'$  (i.e.  $u$  and  $v$  appear in the same subset of leaves of  $F$ ). For all  $i \in \{1, \dots, t\}$ , let  $\mathcal{L}^i \subseteq \mathcal{L}'$  denote the subset of leaves in which vertices of  $C_i$  were before the partitioning. Then, replace the leaves of  $F$  by  $C_1, \dots, C_t$ , and for all  $i \in \{1, \dots, t\}$ , guess  $w(C_i)$  and set  $g(C_i) = \bigcup_{L \in \mathcal{L}^i} g(L)$ .

Let us give the intuition of *Branching Rule 3*. This rule ensures that the set of leaves attached to a same node are vertex disjoint and that the partition was made in such a way that two vertices in the same leaf after the application of the rule were in the same subset of leaves before it. Notice that the remaining Branching Rules can create bad fathers, but decrease  $k$ .

**Lemma 5.** *Branching Rule 3 is safe.*

*Proof.* Notice first that by construction  $\#BF$  decreases, whereas  $k$  and  $\#AL$  remain unchanged.

Let us now check the invariants. Since vertices which appear in a leaf before the transformation still appear on some leaves, invariant 2 is preserved. By Remark 2, no leaf contains a lonely vertex. Thus, all vertices of  $C$  are contained in  $F$  and thus induce a clique. Since we do not modify  $F$ , no vertex nor edge has been removed from the graph, and invariant 1 still holds. For proving that invariant 3 still holds, let  $i \in \{1, \dots, t\}$  and  $u \in C_i$ . Before the partitioning we had:

$$\begin{aligned} N_0(u) \cap S &= \bigcup_{L \in \mathcal{X}_u \cap \mathcal{L}} g(L) \\ &= \left( \bigcup_{L \in \mathcal{X}_u \cap \mathcal{L}'} g(L) \right) \cup \left( \bigcup_{L \in \mathcal{X}_u \cap (\mathcal{L} \setminus \mathcal{L}')} g(L) \right) \end{aligned}$$

And by definition, we now have  $\bigcup_{L \in \mathcal{X}_u \cap \mathcal{L}'} g(L) = g(C_i)$ . Hence, the invariant is preserved.

Let us now turn to invariant 4. Consider a solution  $S^*$  optimal and closed by inclusion satisfying  $S \subseteq S^*$  and the flags  $w$  on the leaves. If we consider the branching where every new leaf  $C_i$  receives the right flag with respect to  $S^* \cap C_i$ , then the solution  $S^*$  satisfies the assigned flags, and invariant 2 holds.  $\square$

**Branching Rule 4: taking a lonely vertex in a father.**

If there exists  $L \in \mathcal{L}$  such that  $\text{pred}(L)$  contains a lonely vertex  $v$ , then take  $v$  in the solution, delete  $k$  by one, and create a new leaf  $N$  adjacent to  $\text{pred}(L)$  and containing vertices of  $L \setminus \{v\}$ . Finally, guess a value for  $w(N)$  and set  $g(N) = \{v\}$ .

**Lemma 6.** *Branching Rule 4 is safe.*

*Proof.* First, it is clear that invariant 1 still holds, since we just removed a vertex  $v$  from the graph, and duplicated a node of the tree in a leaf. Then, since we created a leaf  $N$  containing all neighbors of  $v$ , and since we assigned a value  $w(N)$  for this new leaf, invariant 2 is preserved. Concerning invariant 3, notice that for all  $u \in \text{pred}(L)$ , its neighborhood in the partial solution after the branching rule ( $N_0(u) \cap S$ ) is exactly the union of its neighborhood in the previous partial solution and  $\{v\}$ . By definition of  $g(N)$ , and since  $u$  now belongs to  $N$ , this proves that the invariant is still true.

Let us now prove that invariant 4 still holds. Let  $S^*$  be an optimal solution closed under inclusion which satisfies  $S \subseteq S^*$  and the already assigned flags  $w$ . If  $S^* \cap \text{pred}(L) \neq \emptyset$  then the result is straightforward since  $v$  is lonely. Otherwise, there are two cases:

- first case: there exists  $L \in \text{pred}(L)$  such that  $S^* \cap L \neq \emptyset$ . In this case, let  $u \in S^* \cap L$ . Since  $L$  does not contain any lonely vertex (see Remark 2),  $S^*$  is actually not closed under inclusion, which proves that this case is impossible.
- second case: for all  $L \in \text{pred}(L)$  we have  $S^* \cap L = \emptyset$ . In this case it means that  $N_0(v) \cap S^* = \emptyset$  and thus we can replace any other vertex of  $S^*$  by  $v$  without increasing its cost.

Finally, it is clear that  $k$  decreases. □

**Branching Rule 5: taking an almost lonely vertex in a leaf.**

If there exists  $L \in \mathcal{L}$  such that  $w(L) = 1$  and  $L$  contains an almost lonely vertex  $v$  (thus contained in  $L$  and  $\text{pred}(L)$ ), then take  $v$  in the solution, decrease  $k$  by one, and create a new leaf  $N$  adjacent to  $\text{pred}(L)$  and containing vertices of  $\text{pred}(L) \setminus \{v\}$ . If  $L$  does not become empty, then guess a new value  $w(L)$ . Finally, guess a value  $w(N)$ , add  $v$  into  $g(L)$ , and set  $g(N) = \{v\}$ .

**Lemma 7.** *Branching Rule 5 is safe.*

*Proof.* Since we just removed a vertex from  $G$  and duplicated a node, creating a leaf, invariant 1 still holds. In addition, neighbors of  $v$  in the remaining graph must appear in the new leaf  $N$ , which receives a flag  $w(N)$ . Hence invariant 2 and 3 also hold (notice that we added  $v$  into  $g(L)$ , and that  $g(N)$  has been set to  $\{v\}$ ). Concerning invariant 4, let  $S^*$  be an optimal solution closed under inclusion, such that  $S \subseteq S^*$ , and respecting the flags  $w$ . Let  $x \in S^* \cap L$  (such a vertex must exist, according to  $w(L)$ ), and suppose that  $v \notin S^*$ . By invariant 2 it holds that  $N_0(v) \cap S \subseteq N_0(x) \cap S$ . Since there is no lonely vertex in  $L$  (cf Remark 2), it holds that  $N_0(v) \cap S^* \cap V \subseteq N_0(v) \cap S^* \cap V$ . Since  $S^* = S \cup (S^* \cap V)$ , this proves that invariant 4 is preserved.

Finally,  $k$  strictly decreases.

**End of the Algorithm.**

**Lemma 8.** *If no rule can be applied then either  $G$  is empty or  $k = 0$ .*

*Proof.* Let us first prove that the depth of  $T$  is at most 1 (that is,  $T$  is a star). Suppose by contradiction that there exists an internal node  $F$  of depth at least 1, *i.e.* at least one leaf is adjacent to  $F$ , and  $F \neq X_r$  (and thus  $\text{pred}(F)$  exists). By *Pre-Processing Rule 2* and *Branching Rule 5*, no leaf of  $F$  has an almost lonely vertex. So every vertex which appears in  $F$  and a leaf of  $F$  also appears in  $\text{pred}(F)$  (since otherwise *Branching Rule 3* would apply). In addition, *Pre-Processing Rule 1* ensures that  $F \not\subseteq \text{pred}(F)$ . Then there exists a vertex  $v$  in  $F$  which is not in  $\text{pred}(F)$ . Hence  $v$  must be a lonely vertex of  $F$  and *Branching Rule 4* can be applied, a contradiction.

So  $T$  is a star rooted on  $X_r$ . Since *Branching Rule 3* cannot be applied, leaves of  $X_r$  are vertex disjoint. So every vertex which appears in a leaf is a lonely or an almost lonely vertex. Let  $L$  be such a leaf. If  $w(L) = 0$ , then *Pre-Processing Rule 2* can be applied. Otherwise *Branching rule 1* or *5* can be applied as long as  $X_r$  has a leaf.

Hence  $G$  is now reduced to a clique. If  $k = 0$  then we already have the solution and can output it. If  $k > 0$ , then since each vertex is a lonely one, *Branching Rule 1* can apply and we can thus choose arbitrarily any remaining vertex.

Thus, the algorithm ends when the graph is empty or when  $k = 0$ . If the graph is empty and  $k > 0$ , then we know that the current branching is not the right one, and then the output does not provide an optimal solution. In the other cases, we compare the costs of all produced solutions (in each branching). Since invariant 4 is preserved in all pre-processing and branching rules, one of the branch of the search tree must provide a solution of optimal cost. Therefore the minimum over all the possible branchings provides a solution with an optimal cost, which finishes the proof.  $\square$

According to the introduction and all the safeness lemmas, the size of the search tree is a function of  $k$ . Then, let us remark that all rules can be applied in *FPT* time. This is clear for *Pre-Processing Rules 1* and *2*, as well as for *Branching Rules 1, 2, 4* and *5*. Concerning *Branching Rule 3*, which consists in partitioning a subset of leaves, it runs in *FPT* time as long as  $|\mathcal{L}|$  is a function of  $k$ . This is obviously the case at the beginning of the algorithm (since  $|\mathcal{L}| < k$ ), and the number of leaves only increase by one in *Branching Rule 4* and *5*, and by a function of the previous number of leaves in *Branching Rule 3*. Since the branching rules are applied at most  $f(k)$  times, we get the desired result.

**Theorem 2.** *There is an FPT algorithm for SPARSEST  $k$ -SUBGRAPH in chordal graphs, parameterized by  $k$ .*

However, the running time of the algorithm may be a tower of 2 of height  $k$ , since *Branching Rule 3* may create  $2^t$  new leaves, where  $t$  is the number of previous leaves of the node  $F$ . Nevertheless, we can slightly modify the algorithm in order to obtain a  $O^*(2^{k^2})$  running time<sup>2</sup>. Indeed, after the application of this rule, all leaves  $L$  such that  $w(L) = 0$  can be gathered into one leaf, since all these vertices are not in the solution. And since all leaves are vertex disjoint, the number of leaves  $L$  such that  $w(L) = 1$  is at most  $k$  (since one vertex of each leaf is in the solution). Hence, the number of leaves of  $F$  after the application of *Branching Rule 3* can actually be bounded by  $k + 1$ . Then, as said previously, the only other branching rules which increase the number of leaves are *Branching Rules 4* and *5*, which both add at most one leaf when they are applied. However, since these branching rules are decreasing  $k$ , the maximum number of leaves of a node  $F$  before the application of *Branching Rule 3* is  $2k$ . Hence, this rule (which upper bounds the running time of the algorithm) runs in time  $O^*(2^{O(2k^2)})$  (we have at most  $2^{2k}$  leaves and we choose at most  $k$  leaves such that  $w(L) = 1$ ). For sake of readability, the presented algorithm does not contain this slight modification.

## 5 Kernel Lower Bound of Sparsest $k$ -Subgraph in Chordal Graphs

*Intuition of the proof.* The following kernel lower bound is obtained using a cross-composition. It is an extension of our previous work [17], showing the  $\mathcal{NP}$ -hardness of SPARSEST  $k$ -SUBGRAPH in chordal graphs. Let us first give the intuition of this result, and then explain the modification we apply which leads to the kernel lower bound. We then explicit the whole construction of the cross-composition and give a formal proof of the result.

The  $\mathcal{NP}$ -hardness proof is a reduction from the classical  $k$ -CLIQUE problem in general graphs and roughly works as follows. Given an input instance  $G = (V, E)$ ,  $k \in \mathbb{N}$  of  $k$ -CLIQUE, we first

<sup>2</sup> The  $O^*(\cdot)$  notation avoids polynomial terms.

build a clique  $A$  representing the vertices of  $G$ . We also represent each edge  $e_j = \{u, v\} \in E$  by a gadget  $F_j$ , and connect the representative vertices of  $u$  and  $v$  in  $A$  to some vertices of  $F_j$  (see the left of Figure 1). The reduction will force the solution to take in  $A$  the representatives of  $(n - k)$  vertices of  $G$  (corresponding to the complement of a solution  $S$  of size  $k$  in  $G$ ), and also to take the same number of vertices among each gadget. The key idea is that the cost of a gadget  $F_j$  increases by one if it is adjacent to one of the selected vertices of  $A$ . Thus, since the goal is to minimize the cost, we will try to maximize the number of gadgets adjacent to the representatives of  $S$  (*i.e.* vertices we did not pick in  $A$ ), the maximum being reached when  $S$  is a clique in  $G$ .

To adapt this reduction into a cross-composition, we add an *instance selector* composed of  $2 \log t$  gadgets adjacent to  $A$  (where  $t$  is the number of input instances of the cross-composition) which encodes the binary representation of each instance index. These gadgets have the same structure as the  $F_j$ . For technical reasons, this instance selector has to be duplicated many times, as well as the clique  $A$  which we must duplicate  $t$  times in order to encode the vertex set of each instance. The right of Figure 1 represents the construction in a simplified way. Let us now define formally the gadgets and state their properties.

*Definition of a gadget.* Let  $T \in \mathbb{N}$  (we will set the value of  $T$  later). The vertex set of each gadget is composed of three sets of  $T$  vertices  $X, Y$  and  $Z$ , with  $X = \{x_1, \dots, x_T\}, Y = \{y_1, \dots, y_T\}$  and  $Z = \{z_1, \dots, z_T\}$ . The set  $X$  induces an independent set, the set  $Z$  induces a clique, and there is a  $(T - 1)$ -clique on  $\{y_2, \dots, y_T\}$ . In addition, for all  $i \in \{1, \dots, T\}$ , we connect  $y_i$  to all vertices of  $Z$  and to  $x_i$ . The left of Figure 1 summarizes the construction.

In the following cross-composition, we will force the solution to take  $2T$  vertices among each gadget  $F$ . It is easy to see that the sparsest  $2T$ -subgraph of  $F$  is composed of the sets  $X$  and  $Z$ , which induces  $\binom{T}{2}$  edges. In addition, if we forbid the set  $Z$  to be in the solution (if the gadget is adjacent to some picked vertices of  $A$ ), then the remaining  $2T$  vertices (namely  $X$  and  $Y$ ) induce  $(\binom{T}{2} + 1)$  edges.

**Theorem 3.** SPARSEST  $k$ -SUBGRAPH does not admit a polynomial kernel in chordal graphs unless  $\mathcal{NP} \subseteq \text{coNP}/\text{poly}$  (parameterized by  $k$ ).

*Proof.* Let  $(G_1, k_1), \dots, (G_t, k_t)$  be a sequence of  $t$  instances of  $k$ -CLIQUE, with  $G_i = (V_i, E_i)$  for all  $i \in \{1, \dots, t\}$ . W.l.o.g. we suppose that  $t = 2^q$  for some  $q \in \mathbb{N}$ , and define  $T = n(n - k)$  and  $M = n^6$ .

Our polynomial equivalence relation is the following: for  $1 \leq i, j \leq t$ ,  $(G_i, k_i)$  is equivalent to  $(G_j, k_j)$  if  $|V_i| = |V_j| = n$ ,  $|E_i| = |E_j| = m$  and  $k_i = k_j = k$ . One can verify that this relation is a polynomial equivalence relation. In what follows we suppose that all instances of the sequence are in the same equivalence class. The output instance  $G' = (V', E'), k', C'$  is defined as follows (see Figure 1):

- For each  $i \in \{1, \dots, t\}$  we construct a clique  $A^i$  on  $n^2$  vertices, where  $A^i$  is composed of  $n$  subcliques  $A_1^i, \dots, A_n^i$ . We also add all possible edges between all cliques  $(A^i)_{i=1..t}$ . Hence,  $A = \bigcup_{i=1}^t A^i$  is a clique of size  $tn^2$ .
- Since all instances have the same number of edges, we construct  $m$  gadgets  $(F_j)_{j=1..m}$ , where each  $F_j$  is composed of  $X_j, Y_j$  and  $Z_j$  as described previously. For all  $i \in \{1, \dots, t\}$ , if there is an edge  $e_j = \{u, v\} \in E_i$ , then we connect all vertices of  $Z_j$  to all vertices of  $A_u^i$  and  $A_v^i$ . Let us define  $\mathcal{F} = \bigcup_{j=1}^m F_j$  the subgraph of all gadgets of the "edge selector".
- We add  $2qM$  gadgets  $(F_{\alpha_j^h})_{j=1..q}^{h=1..M}$  and  $(F_{\beta_j^h})_{j=1..q}^{h=1..M}$ , where all gadgets are isomorphic to the edge gadgets, and thus composed of  $X_{\alpha_j^h}, Y_{\alpha_j^h}$  and  $Z_{\alpha_j^h}$  (resp.  $X_{\beta_j^h}, Y_{\beta_j^h}$  and  $Z_{\beta_j^h}$ ) for all  $h \in \{1, \dots, M\}$  and all  $j \in \{1, \dots, q\}$ . Let  $i \in \{1, \dots, t\}$ , and consider its binary representation  $b \in \{0, 1\}^q$ . For

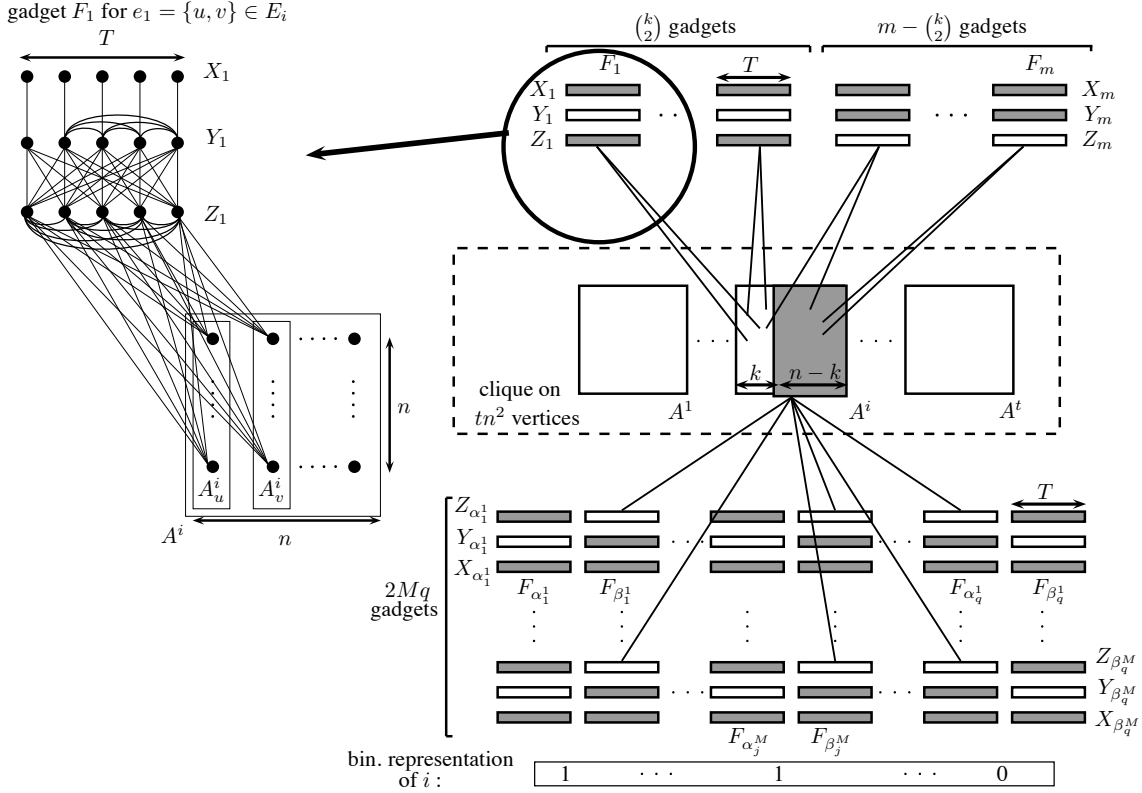


Fig. 1: Schema of the cross-composition (right) and a detailed gadget (left). Grey rectangles represent vertices of the solution, supposing that  $G_i$  contains a clique of size  $k$ . Notice that gadgets of the bottom have been drawn in the reverse direction (e.g.  $X_{\beta_1^1}$  is below  $Y_{\beta_1^1}$ ). Edges of the clique  $A$  have not been drawn for sake of clarity.

- all  $j \in \{1, \dots, q\}$ , if the  $j^{\text{th}}$  bit of  $b$  equals 0, then connect all vertices of  $A^i$  to all vertices of  $\bigcup_{h=1}^M Z_{\alpha_j^h}$ . Otherwise, connect all vertices of  $A^i$  to all vertices of  $\bigcup_{h=1}^M Z_{\beta_j^h}$ . Let us define  $\mathcal{B} = \bigcup_{h=1}^M \bigcup_{j=1}^q (F_{\alpha_j^h} \cup F_{\beta_j^h})$  the subgraph of all gadgets of the "instance selector".
- We set  $k' = T + 2Tm + 4TqM$  and  $C' = \binom{T}{2} + \binom{T}{2}(m + 2Mq) + (m - \binom{k}{2}) + Mq$ .

It is clear that  $G', k'$  and  $C'$  can be constructed in time polynomial in  $\sum_{i=1}^t |G_i| + k_i$ . Then, one can verify that  $G'$  is a chordal graph. Indeed, it is known [12] that a graph is chordal if and only if one can repeatedly find a simplicial vertex (a vertex whose neighborhood is a clique) and delete it from the graph until it becomes empty. Such an ordering is called a simplicial elimination order. It is easily seen that for each gadget,  $X, Y$  and then  $Z$  is a simplicial elimination order (each gadget is only adjacent to the clique  $A$  via its set  $Z$ ). Finally it remains the clique  $A$  which can be eliminated.

In addition, notice that the parameter  $k'$  is a polynomial in  $n, k$  and  $\log t$  only and thus respect the definition of a cross-composition. We finally prove that there exists  $i \in \{1, \dots, t\}$  such that  $G_i$

contains a clique  $K$  of size  $k$  if and only if  $G'$  contains a set  $K'$  of  $k'$  vertices inducing  $C'$  edges or less.

**Lemma 9.** *If there exists  $i \in \{1, \dots, t\}$  such that  $G_i$  contains a  $k$ -clique, then  $G'$  contains  $k'$  vertices inducing at most  $C'$  edges.*

*Proof.* Suppose that  $K \subseteq V_i$  is a clique of size  $k$  in  $G_i$ . W.l.o.g. suppose that  $K = \{v_1, \dots, v_k\}$ , and that  $\{\{u, v\}, u, v \in K\} = \{e_1, \dots, e_{\binom{k}{2}}\}$ . Let  $b \in \{0, 1\}^q$  be the binary representation of  $i$ . We build  $K'$  as follows (see Figure 1).

- For all  $j \in \{1, \dots, \binom{k}{2}\}$ ,  $K'$  contains  $X_j$  and  $Z_j$  ( $2T$  vertices inducing  $\binom{T}{2}$  edges for each gadget  $F_j$ ).
- For all  $j \in \{\binom{k}{2} + 1, \dots, m\}$ ,  $K'$  contains  $X_j$  and  $Y_j$ . ( $2T$  vertices inducing  $(\binom{T}{2} + 1)$  edges for each gadget  $F_j$ ).
- For all  $u \notin \{1, \dots, k\}$ ,  $K'$  contains  $A_u^i$  ( $T$  vertices inducing  $\binom{T}{2}$  edges).
- For all  $h \in \{1, \dots, M\}$ , and all  $j \in \{1, \dots, q\}$ ,  $K'$  contains  $X_{\alpha_j^h}$  and  $X_{\beta_j^h}$ . Moreover, if the  $j^{\text{th}}$  bit of  $b$  equals 1, then  $K'$  contains  $Y_{\beta_j^h}$  and  $Z_{\alpha_j^h}$ , otherwise  $K'$  contains  $Z_{\beta_j^h}$  and  $Y_{\alpha_j^h}$  ( $4T$  vertices inducing  $(2\binom{T}{2} + 1)$  edges for each pair of gadgets  $F_{\alpha_j^h}$  and  $F_{\beta_j^h}$ ).

One can easily verify that  $K'$  is a set of  $k'$  vertices inducing  $C'$  edges.  $\square$

The following lemma terminates the proof.

**Lemma 10.** *If  $G'$  contains  $k'$  vertices inducing at most  $C'$  edges, then  $\exists i \in \{1, \dots, t\}$  such that  $G_i$  contains a  $k$ -clique.*

*Proof.* Suppose now that  $K'$  is a set of  $k'$  vertices inducing  $C'$  edges. For a set  $S \subseteq V'$ , we denote by  $tr(S) = S \cap K'$  the trace of the solution on  $S$ . For all  $v \in V'$ , let  $\mu(v) = |tr(N(v))|$  be the number of neighbors of  $v$  belonging to  $K'$ .

Let  $\mathcal{I} = \{1, \dots, m\} \cup \{\alpha_j^h\}_{j=1..q}^{h=1..M} \cup \{\beta_j^h\}_{j=1..q}^{h=1..M}$  be the set of all indices of gadgets. As in the definition of the gadgets given above, we define for all  $\gamma \in \mathcal{I}$  the sets  $X_\gamma = \{x_1^\gamma, \dots, x_T^\gamma\}$ ,  $Y_\gamma = \{y_1^\gamma, \dots, y_T^\gamma\}$  and  $Z_\gamma = \{z_1^\gamma, \dots, z_T^\gamma\}$ .

We define  $E_0 = \{\gamma \in \mathcal{I} \text{ such that } \forall x \in tr(A), \text{ no vertex of } Z_\gamma \text{ is adjacent to } x\}$ , *i.e.*  $E_0$  represents the indices of all gadgets  $F_\gamma$  which are not adjacent to vertices of the solution among the clique  $A$ . Then, define  $E_1 = \mathcal{I} \setminus E_0$ , which represents indices of gadgets which are adjacent to at least one vertex of  $tr(A)$ .

In the three following Claims (1, 2 and 3), we show that we can restructure the solution inside each gadget in order to encode a solution for the  $k$ -clique instance. To do so, we define the notion of *safe replacement*:

Let  $u \in K'$  and  $v \in V' \setminus K'$ . We say that  $(K' \setminus \{u\}) \cup \{v\}$  is a safe replacement if we have  $\mu(v) \leq \mu(u)$  if  $\{u, v\} \notin E'$  and  $\mu(v) - 1 \leq \mu(u)$  if  $\{u, v\} \in E'$ . It is easily seen that in this case  $(K' \setminus \{u\}) \cup \{v\}$  does not induce more edges than  $K'$ . For the sake of readability, we will keep the same notations and update the set  $K'$  when applying replacements, as well as the sets  $E_0$  and  $E_1$  when replacing vertices of  $A$  (*e.g.* if there exists  $\gamma \in E_1$  such that  $F_\gamma$  is adjacent to a unique vertex  $u \in tr(A)$ , and if a replacement removes  $u$  from the solution, then  $\gamma$  now belongs to  $E_0$ ).

**Claim 1** *Without loss of generality (and optimality of  $K'$ ), we can suppose that for all  $\gamma \in \mathcal{I}$  we have  $X_\gamma \subseteq K'$ .*

*Proof.* Let  $S = \bigcup_{\gamma \in \mathcal{I}} X_\gamma$ . Since we have  $k' > |S|$ , we have  $K' \setminus S \neq \emptyset$ . Suppose that there exists  $\gamma \in \mathcal{I}$  and  $i \in \{1, \dots, T\}$  such that  $x_i^\gamma \notin K'$ . Recall that  $y_i^\gamma$  is the only neighbor of  $x_i^\gamma$ . If  $y_i^\gamma \notin K'$ , then we have  $\mu(x_i^\gamma) = 0$  and we can thus safely replace any other vertex of  $K' \setminus S$  by  $x_i^\gamma$ . Now, if  $y_i^\gamma \in K'$ , then  $\mu(x_i^\gamma) = 1$ . Since  $x_i^\gamma$  and  $y_i^\gamma$  are adjacent,  $(K' \setminus \{y_i^\gamma\}) \cup \{x_i^\gamma\}$  is a safe replacement.  $\square$

In the following, we suppose that for all  $\gamma \in \mathcal{I}$  we have  $X_\gamma \subseteq K'$ .

**Claim 2**  $K'$  can be safely modified such that one of the two following cases must happen (see Figure 2):

- case A1: for all  $\gamma \in E_0$  we have  $tr(Z_\gamma) = Z_\gamma$ .
- case A2: for all  $\gamma \in E_0$  we have  $tr(Y_\gamma) = \emptyset$ .

*Proof.* Let us first restructure each gadget of  $E_0$  separately. For all  $\gamma \in E_0$  such that  $tr(Y_\gamma) \neq \emptyset$  and  $tr(Z_\gamma) \neq Z_\gamma$ , let  $j_0 = \max\{j \in \{1, \dots, T\} : y_j^\gamma \in tr(Y_\gamma)\}$  and let  $j_1$  be such that  $z_{j_1}^\gamma \notin tr(Z_\gamma)$ . Recall that Claim 1 ensures that  $x_{j_0}^\gamma$  is in  $K'$ . If  $j_0 \neq 1$ , then  $\mu(y_{j_0}^\gamma) = y + z + 1$ , where  $y = |N(y_{j_0}^\gamma) \cap tr(Y_\gamma)|$  and  $z = |N(y_{j_0}^\gamma) \cap tr(Z_\gamma)|$ . On the other side, we have  $\mu(z_{j_1}^\gamma) \leq y + z + 1$  (more precisely,  $\mu(z_{j_1}^\gamma) = y + z + 1$  if  $y_1^\gamma \in K'$ , and  $\mu(z_{j_1}^\gamma) = y + z$  if  $y_1^\gamma \notin K'$ ). Roughly speaking, this switch ensures that we necessarily “loose” the edge due to the vertex of  $X_\gamma$  and we gain at most one edge due to  $y_1^\gamma$ . Hence  $\mu(z_{j_1}^\gamma) \leq \mu(y_{j_0}^\gamma)$  and  $(K' \setminus \{y_{j_0}^\gamma\}) \cup \{z_{j_1}^\gamma\}$  is a safe replacement. If  $j_0 = 1$ , then it means that  $tr(Y_\gamma) = \{y_1^\gamma\}$ . Suppose that there exists  $j_1$  such that  $z_{j_1}^\gamma \notin tr(Z_\gamma)$ . We have  $\mu(y_1^\gamma) = z + 1$  where  $z = |N(y_1^\gamma) \cap tr(Z_\gamma)|$ , and  $\mu(z_{j_1}^\gamma) = z + 1$ . Here again  $(K' \setminus \{y_1^\gamma\}) \cup \{z_{j_1}^\gamma\}$  is a safe replacement. After all these replacements, given any  $\gamma \in E_0$ ,  $tr(Y_\gamma) \neq \emptyset$  implies that  $tr(Z_\gamma) = Z_\gamma$ .

Then, we proceed to replacements between gadgets  $F_\gamma$ ,  $\gamma \in E_0$ . If one can find  $a, b \in E_0$  such that  $tr(Y_a) \neq \emptyset$  and  $tr(Z_b) \neq Z_b$ , then let  $j_0$  be such that  $y_{j_0}^a \in tr(Y_a)$  and let  $j_1$  be such that  $z_{j_1}^b \notin tr(Z_b)$ . We have  $\mu(y_{j_0}^a) \geq T + 1$  and  $\mu(z_{j_1}^b) \leq T - 1$ . Thus,  $(K' \setminus \{y_{j_0}^a\}) \cup \{z_{j_1}^b\}$  is a safe replacement.

These replacements end either when  $tr(Y_\gamma) = \emptyset$  for all  $\gamma \in E_0$  or when  $tr(Z_\gamma) = Z_\gamma$  for all  $\gamma \in E_0$ , which achieves the proof of Claim 1.  $\square$

**Claim 3**  $K'$  can be safely modified such that one of the two following cases must happen (see Figure 2):

- case B1: for all  $\gamma \in E_1$  we have  $tr(Y_\gamma) = Y_\gamma$ .
- case B2: for all  $\gamma \in E_1$  we have  $tr(Z_\gamma) = \emptyset$ .

*Proof.* The proof is roughly based on the fact that replacing a vertex of  $Z_\gamma$  by a vertex of  $Y_\gamma$  permits to “loose” at least one edge with vertices  $A$  and “gain” one edge with a vertex of  $X_\gamma$ . Let us formally prove Claim 3. Similarly to the proof of Claim 2, we first restructure each gadget of  $E_1$  separately: for all  $\gamma \in E_1$  such that  $tr(Z_\gamma) \neq \emptyset$  and  $tr(Y_\gamma) \neq Y_\gamma$ , let  $j_0 = \max\{j \in \{1, \dots, T\} : y_j^\gamma \notin K'\}$  and let  $j_1$  be such that  $z_{j_1}^\gamma \in tr(Z_\gamma)$ . Recall that by definition of  $E_1$ , there exists  $i, j \in \{1, \dots, n\}$  such that  $z_{j_1}^\gamma$  is adjacent to  $a_i^j$ . We have  $\mu(z_{j_1}^\gamma) \geq y + z + 1$ , where  $y = |N(z_{j_1}^\gamma) \cap Y_\gamma|$  and  $z = |N(z_{j_1}^\gamma) \cap Z_\gamma|$ . On the other side, we have  $\mu(y_{j_0}^\gamma) \leq z + y + 2$  (indeed,  $|N(y_{j_0}^\gamma) \cap Z_\gamma| = z + 1$ ,  $|N(y_{j_0}^\gamma) \cap Y_\gamma| \leq y$  and  $|N(y_{j_0}^\gamma) \cap X_\gamma| = 1$ ). Since  $\{y_{j_0}^\gamma, z_{j_1}^\gamma\} \in E'$ , it holds that  $(K' \setminus \{z_{j_1}^\gamma\}) \cup \{y_{j_0}^\gamma\}$  is a safe replacement. After all these replacements, given any  $\gamma \in E_1$ ,  $tr(Z_\gamma) \neq \emptyset$  implies that  $tr(Y_\gamma) = Y_\gamma$ .

We now proceed to replacements between gadgets  $F_\gamma$ ,  $\gamma \in E_1$ . If one can find  $a, b \in E_1$  such that  $tr(Z_a) \neq \emptyset$  and  $tr(Y_b) \neq Y_b$ , then let  $j_0$  be such that  $y_{j_0}^b \notin tr(Y_b)$  and let  $j_1$  be such that  $z_{j_1}^a \in tr(Z_a)$ . We have  $\mu(z_{j_1}^a) \geq T + 1$  and  $\mu(y_{j_0}^b) \leq T - 1$ . Thus  $(K' \setminus \{z_{j_1}^a\}) \cup \{y_{j_0}^b\}$  is a safe replacement.

As previously, the replacements ends either when  $tr(Y_\gamma) = Y_\gamma$  for all  $\gamma \in E_1$  or when  $tr(Z_\gamma) = \emptyset$  for all  $\gamma \in E_1$ .  $\square$

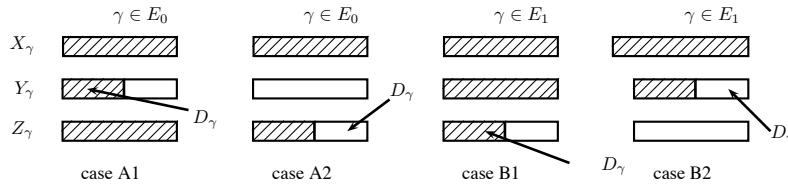


Fig. 2: Schema of different cases. Shaded rectangles represent part of  $K'$ .

We now define for each case and each  $\gamma \in \mathcal{I}$  the set of vertices  $D_\gamma \subseteq Y_\gamma \cup Z_\gamma$  that have to be replaced:

- case A1: for all  $\gamma \in E_0$ ,  $D_\gamma = Y_\gamma \cap K'$
- case A2: for all  $\gamma \in E_0$ ,  $D_\gamma = Z_\gamma \setminus K'$
- case B1: for all  $\gamma \in E_1$ ,  $D_\gamma = Z_\gamma \cap K'$
- case B2: for all  $\gamma \in E_1$ ,  $D_\gamma = Y_\gamma \setminus K'$

Notice that if  $D_\gamma = \emptyset$  for all  $\gamma \in E_0$  (resp. for all  $\gamma \in E_1$ ), then cases A1 and A2 (resp. B1 and B2) collapse. If such a case happen for all  $\gamma \in \mathcal{I}$ , we can immediately conclude, as we will see in Claim 5. Now, we will show that if cases A1 and B1 happen (or if  $D_\gamma = \emptyset$  for all  $\gamma \in \mathcal{I}$ ), then the solution must hit the clique  $A$  in only one subclique  $A^i$  for some  $i \in \{1, \dots, t\}$ :

**Claim 4** *If cases A1 and B1 happen (or if  $D_\gamma = \emptyset$  for all  $\gamma \in \mathcal{I}$ ), then there exists  $i \in \{1, \dots, t\}$  such that  $\text{tr}(A) \subseteq A^i$ , i.e. the solution  $K'$  only appears in one clique  $A^i$  among  $A$ .*

*Proof.* Let  $\Delta = \sum_{\gamma \in \mathcal{I}} |D_\gamma|$ , and suppose by contradiction that there exists  $i, j \in \{1, \dots, t\}$  with  $i \neq j$  such that  $K' \cap A^i \neq \emptyset$  and  $K' \cap A^j \neq \emptyset$ . First, since we are in case A1 and B1, the number of edges induced by each gadget is at least  $\binom{T}{2}$ . Then, let  $S$  (resp.  $\bar{S}$ ) be the number of pairs of gadgets corresponding to a bit on which the binary representations of  $i$  and  $j$  is the same (resp. differ). Recall that  $S + \bar{S} = Mq$ . Then, since  $i \neq j$ , the binary representations of  $i$  and  $j$  must differ on at least one bit, which implies  $\bar{S} \geq M$ . Let us count the number of edges induced by each pair of gadget, whether they correspond to a bit value shared by the binary representation of  $i$  and  $j$  or not.

Let  $p \in \{1, \dots, q\}$  such that the binary representations of  $i$  and  $j$  are the same. Then, for all  $h \in \{1, \dots, M\}$ , three cases may happen:

- $Y_{\alpha_p^h} \subseteq K'$  and  $Y_{\beta_p^h} \subseteq K'$ . In this case the pair of gadgets induces at least  $2\binom{T}{2} + 2$  edges.
- $Y_{\alpha_p^h} \subseteq K'$  and  $Z_{\beta_p^h} \subseteq K'$  (or the contrary). In this case the pair of gadgets induces at least  $2\binom{T}{2} + 1$  edges.
- $Z_{\alpha_p^h} \subseteq K'$  and  $Z_{\beta_p^h} \subseteq K'$ . In this case the pair of gadgets induces at least  $2\binom{T}{2} + T$  edges, since either  $Z_{\alpha_p^h}$  or  $Z_{\beta_p^h}$  is adjacent to at least one vertex of  $\text{tr}(A^i)$ .

Hence, in all three cases the solution in each pair of such gadgets induces at least  $(2\binom{T}{2} + 1)$  edges.

Let us now focus on some  $p \in \{1, \dots, q\}$  such that the binary representations of  $i$  and  $j$  differ. Then, for all  $h \in \{1, \dots, M\}$ , notice that both  $Z_{\alpha_p^h}$  and  $Z_{\beta_p^h}$  are adjacent to at least one vertex in  $A^i \cup A_j$ . Here again three cases may happen:



- $Y_{\alpha_p^h} \subseteq K'$  and  $Y_{\beta_p^h} \subseteq K'$ . In this case the pair of gadgets induces at least  $2\binom{T}{2} + 2$  edges.
- $Y_{\alpha_p^h} \subseteq K'$  and  $Z_{\beta_p^h} \subseteq K'$  (or the contrary). In this case the pair of gadgets induces at least  $2\binom{T}{2} + T + 1$  edges.
- $Z_{\alpha_p^h} \subseteq K'$  and  $Z_{\beta_p^h} \subseteq K'$ . In this case the pair of gadgets induces at least  $2\binom{T}{2} + 2T$  edges.

Hence, in all three cases the solution in each pair of such gadgets induces at least  $(2\binom{T}{2} + 2)$  edges. In addition, it is easily seen that the number of edges induced by  $tr(A)$  is  $\binom{T}{2} - \binom{\Delta}{2} - \Delta(T - \Delta)$ , since it is a clique of size  $(T - \Delta)$ . To resume:

- $tr(A)$  induces  $(\binom{T}{2} - \binom{\Delta}{2} - \Delta(T - \Delta))$  edges.
- Each gadget (both from the edge or the instance selector) induces at least  $\binom{T}{2}$  edges (there are  $(m + 2Mq)$  gadgets), and more precisely:
  - Each pair of gadgets corresponding to a shared bit value of the binary representation of  $i$  and  $j$  induces  $(\binom{T}{2} + 1)$  edge (*i.e.* one more than the "normal" ones). There are  $S$  such pairs of gadgets.
  - Each pair of gadgets corresponding to a different bit value of the binary representation of  $i$  and  $j$  induces  $(\binom{T}{2} + 2)$  edge (*i.e.* two more than the "normal" ones). There are  $\bar{S}$  such pairs of gadgets.

Thus we have:

$$\begin{aligned}
E(K') &\geq \binom{T}{2} - \binom{\Delta}{2} - \Delta(T - \Delta) + \binom{T}{2}(m + 2Mq) + S + 2\bar{S} \\
&= \binom{T}{2} - \binom{\Delta}{2} - \Delta(T - \Delta) + \binom{T}{2}(m + 2Mq) + Mq + \bar{S} \\
&\geq \binom{T}{2} - \binom{\Delta}{2} - \Delta(T - \Delta) + \binom{T}{2}(m + 2Mq) + Mq + M
\end{aligned}$$

And thus

$$E(K') - C' \geq M - m + \binom{k}{2} - \binom{\Delta}{2} - \Delta(T - \Delta)$$

Since  $M = n^6$ , we have  $E(K') > C'$  which is impossible.  $\square$

**Claim 5** *If  $D_\gamma = \emptyset$  for all  $\gamma \in \mathcal{I}$ , then there exists  $i \in \{1, \dots, t\}$  such that  $G$  contains a clique of size  $k$ .*

*Proof.* By construction, we have  $|tr(A)| = T$  and  $|tr(F_\gamma)| = 2T$  for all  $\gamma \in \mathcal{I}$ . Thus,  $E(tr(A)) = \binom{T}{2}$  and  $E(tr(F_\gamma)) = \binom{T}{2} + 1$  if  $\gamma \in E_1$ , and  $E(tr(F_e)) = \binom{T}{2}$  if  $\gamma \in E_0$ . Hence, we have  $E(K') \geq \binom{T}{2} + \binom{T}{2}(m + 2Mq) + |E_1|$ .

By Claim 4, there exists  $i \in \{1, \dots, t\}$  such that  $tr(A) \subseteq A^i$ . Thus, there are at most  $Mq$  gadgets among the instance selector which are not adjacent to  $tr(A)$ , and which can belong to  $E_0$ . This implies that there are at least  $Mq$  gadgets among the instance selector which must belong to  $E_1$ . Let  $E_0^e = E_0 \cap \{1, \dots, m\}$  be the restriction of  $E_0$  in the edge selector, and similarly  $E_1^e = \{1, \dots, m\} \setminus E_0^e$ . The arguments above show that  $|E_1^e| \leq m - \binom{k}{2}$ , which implies  $|E_0^e| \geq \binom{k}{2}$ . In addition, each gadget  $j \in E_0^e$  corresponding to the edge  $e_j = \{u, v\}$  of  $G_i$  is adjacent to the cliques  $A_u^i$  and  $A_v^i$ , which must be such that  $A_u^i \cap K' = \emptyset$  and  $A_v^i \cap K' = \emptyset$  by definition of  $E_0$ . However, since  $|tr(A)| = |tr(A^i)| = T$ , the number of such cliques is at most  $n - \lfloor \frac{T}{n} \rfloor = k$ . This proves that these  $|E_0^e|$  edges of  $G$  can be induced by at most  $k$  vertices, *i.e.*  $G_i$  contains a clique of size  $k$ .  $\square$

Let us now combine the four possible cases of Claims 2 and 3:

- Case A1 and B1: let  $\Delta = \sum_{\gamma \in \mathcal{I}} |D_\gamma|$ , and suppose that  $\Delta > 0$  (otherwise we conclude by Claim 5). Let us count the number of edges induced by such a solution. To do so, we count the number of edges induced by the solution among vertices of  $A$ , and the number of edges covered by the solution among the gadgets. First, it is clear that  $|tr(A)| = T - \Delta$ , and thus the number of edges induced by  $tr(A)$  is  $\binom{T}{2} - \binom{\Delta}{2} - \Delta(T - \Delta)$  since  $A$  is a clique. In addition, since we are in case A1 and B1, the trace of the solution in all gadgets (both from the edge or the instance selector) covers at least  $\binom{T}{2}$  edges. More precisely, for each gadget  $\gamma \in \mathcal{I}$  three cases may happen:
  - if  $D_\gamma = \emptyset$ , then  $tr(F_\gamma)$  covers exactly  $\binom{T}{2}$  edges if  $\gamma \in E_0$  and exactly  $\binom{T}{2}$  edges if  $\gamma \in E_1$ .
  - if  $D_\gamma \neq \emptyset$ , then:
    - \* if  $\gamma \in E_0$ , then since each vertex of  $Y_\gamma$  is connected to all vertices of  $Z_\gamma$  and to one vertex of  $X_\gamma$ , we have that  $tr(F_\gamma)$  covers exactly  $(\binom{T}{2} + |D_\gamma|(T + 1))$  edges (see Figure 2).
    - \* if  $\gamma \in E_1$ , then since each vertex of  $Z_\gamma$  is connected to all vertices of  $Y_\gamma$ , and to at least one vertex of  $tr(A)$ , we have that  $tr(F_\gamma)$  covers at least  $(\binom{T}{2} + 1 + |D_\gamma|(T + 1))$  edges (recall that if  $D_\gamma$  the gadgets covers exactly  $(\binom{T}{2} + 1)$  edges).

Summing up to all gadgets, the solution among all gadgets covers  $(\binom{T}{2})(m + 2Mq) + |E_1| + \Delta(T + 1)$  edges.

We define  $E_0^e = E_0 \cap \{1, \dots, m\}$  the restriction of  $E_0$  to the edge selector and  $E_0^b = E_0 \setminus E_0^e$  the restriction of  $E_0$  to the instance selector, as well as the corresponding sets  $E_1^e = E_1 \cap \{1, \dots, m\}$  and  $E_1^b = E_1 \setminus E_1^e$ .

By Claim 4, there exist  $i \in \{1, \dots, t\}$  such that  $tr(A) \subseteq A^i$ . This implies that  $|E_0^b| = |E_1^b| = Mq$  (roughly speaking, for each pair of gadgets of the instance selector, only one of the two is connected to  $A_i$  and thus to  $tr(A)$ , depending on the corresponding bit value). Thus, we have  $|E_1| = Mq + |E_1^e| = Mq + m - |E_0^e|$ . Combining all these, we obtain:

$$E(K') \geq \binom{T}{2} - \binom{\Delta}{2} - \Delta(T - \Delta) + \binom{T}{2}(m + 2Mq) + \Delta(T + 1) + Mq + m - |E_0^e|$$

And thus:

$$\begin{aligned} E(K') - C' &\geq \binom{k}{2} + \Delta(T + 1) - |E_0^e| - \binom{\Delta}{2} - \Delta(T - \Delta) \\ &= \frac{\Delta(\Delta + 3)}{2} + \binom{k}{2} - |E_0^e| \end{aligned}$$

Thus, since we supposed  $E(K') - C' \leq 0$  it implies

$$|E_0^e| \geq \frac{\Delta(\Delta + 3)}{2} + \binom{k}{2} \tag{1}$$

On the other hand, the number of vertices of  $G_i$  inducing all edges of  $E_0^e$  is at most  $k + \lfloor \frac{\Delta}{n} \rfloor$ . Hence we have  $|E_0^e| \leq \binom{k + \lfloor \frac{\Delta}{n} \rfloor}{2}$ . Hence we have:

$$\binom{k + \lfloor \frac{\Delta}{n} \rfloor}{2} \geq \frac{\Delta(\Delta + 3)}{2} + \binom{k}{2}$$

If  $\Delta < n$ , then  $\lfloor \frac{\Delta}{n} \rfloor = 0$  and it contradicts the previous inequality. If  $\Delta \geq n$ , then it contradicts inequality (1) since we have by definition  $|E_0^e| \leq m$ . Hence, we must have  $\Delta = 0$  and the result follows by Claim 5.

- Case A2 and B2: let  $\Delta_0 = \sum_{\gamma \in E_0} |D_\gamma|$ ,  $\Delta_1 = \sum_{\gamma \in E_1} |D_\gamma|$ , and  $\Delta = \Delta_0 + \Delta_1$ , and suppose that  $\Delta > 0$  (otherwise we conclude by Claim 5). Let us notice that for all  $u \in tr(A)$ ,  $\mu(u) \geq T$ . On the other hand, for all  $\gamma \in \mathcal{I}$  such that there exists  $v \in D_\gamma$ , we have  $\mu(v) \leq T$  (remark that if  $\gamma \in E_1$ , then  $D_\gamma \subseteq Y_\gamma$ , and if  $\gamma \in E_0$ , then  $v$  is not adjacent to  $tr(A)$  by definition of  $E_0$ ). Thus  $(K' \setminus \{u\}) \cup \{v\}$  is a safe replacement. Since before this replacement we had  $tr(A) = T + \Delta$ , it is clear that we can repeat this replacement (*i.e.*  $(K' \setminus \{u\}) \cup \{v\}$  where  $u \in tr(A)$  and  $v \in D_\gamma$  for some  $\gamma \in \mathcal{I}$ )  $\Delta$  times safely. At this point, the updated value of  $\Delta$  is 0, *i.e.*  $D_\gamma = \emptyset$  for all  $\gamma \in \mathcal{I}$ . We then conclude by Claim 5.
- Case A2 and B1: if there exists  $\gamma \in E_0$  such that there exists  $u \in D_\gamma$ , then  $\mu(u) < T$ . If such a vertex exists, then either  $|tr(A)| > T$  or there exists  $\gamma' \in E_1$  such that there exists  $v \in D_{\gamma'}$ . In the first case for all  $x \in tr(A)$  we have  $\mu(x) \geq T$ , and  $(K' \setminus \{x\}) \cup \{u\}$  is a safe replacement. In the second case we have  $\mu(v) > T$  and here again  $(K' \setminus \{v\}) \cup \{u\}$  is a safe replacement. After these replacements we must have  $D_\gamma = \emptyset$  for all  $\gamma \in E_0$ , and we can apply the case A1 and B1.
- Case A1 and B2: if there exists  $\gamma \in E_1$  such that there exists  $u \in D_\gamma$ , then  $\mu(u) < T$ . If such a vertex exists, then either  $|tr(A)| > T$  or there exists  $\gamma' \in E_0$  such that there exists  $v \in D_{\gamma'}$ . In the first case for all  $x \in tr(A)$  we have  $\mu(x) \geq T$ , and  $(K' \setminus \{x\}) \cup \{u\}$  is a safe replacement. In the second case we have  $\mu(v) > T$  and here again  $(K' \setminus \{v\}) \cup \{u\}$  is a safe replacement. After these replacements we must have  $D_\gamma = \emptyset$  for all  $\gamma \in E_1$ , and we can apply the case A1 and B1.

□

## 6 Dynamic Programming in Interval Graphs for Sparsest $k$ -Subgraph

### 6.1 Introduction

In this section we provide a dynamic programming for SPARSEST  $k$ -SUBGRAPH in interval graphs, leading to an *FPT* algorithm parameterized by  $C$ , the number of edges in the solution. Remark that the parameterization by  $C$  is stronger than the natural parameterization by  $k$ , as we always have  $C \leq \binom{k}{2}$ . In addition, for  $C = 0$ , we obtain the known polynomial algorithm for INDEPENDENT SET in interval graphs. Notice that the complexity ( $\mathcal{NP}$ -hard versus *Polynomial*) is still unknown for SPARSEST  $k$ -SUBGRAPH in interval graphs. However, similarly to the long-standing open problem of DENSEST  $k$ -SUBGRAPH in the same graph class [8], we believe that is may be a tough question.

The algorithm works as follows: we first sort intervals according to their right endpoints. Then we process the graph from the left to the right, and construct all possible connected components composed of at most  $k$  vertices and inducing less than  $C$  edges, and make a recursive call at the right of this constructed connected component. Using swapping arguments, we show that we only have to enumerate an *FPT* number of such connected component at each step.

## 6.2 Preliminaries

In the following, we will make no distinction between a vertex and its corresponding interval, as well as we will make no distinction between edges in the graph and overlaps in the corresponding interval model.

For the rest of the section,  $G = (V, E)$  will denote the input graph of the problem. The associated interval set will be denoted by  $\mathcal{I} = \{I_1, \dots, I_n\}$ . Without loss of polynomiality, we suppose that all endpoints are pairwise distinct. Given  $I \in \mathcal{I}$ , we denote by  $right(I) \in \mathbb{R}$  (resp.  $left(I) \in \mathbb{R}$ ) the right (resp. left) endpoint of  $I$ . By extension, for any set  $S \subseteq \mathcal{I}$ , we define  $left(S) = \arg \min_{I \in S} left(I)$  (resp.  $right(S) = \arg \max_{I \in S} right(I)$ ). Unless otherwise stated, we suppose that  $\mathcal{I}$  is sorted according to the right endpoints of the intervals (*i.e.* for all  $i \in \{2, \dots, n\}$  we have  $right(I_{i-1}) < right(I_i)$ ). For  $S \subseteq \mathcal{I}$  and  $r \leq |S|$ , we define the " $r$ -leftmost intervals of  $S$ " as the  $r$  first intervals in an ordering of  $S$  (where intervals are sorted according to their right endpoints). Notice that the 1-leftmost interval will simply be called *the leftmost interval*. Given a set  $S \subseteq \mathcal{I}$ , we denote  $cost(S)$  the number of edges in the graph induced by intervals of  $S$ .

Given  $x \in \mathbb{R}$  we define  $\mathcal{I}_{\geq x} = \{I \in \mathcal{I} : x \leq left(I)\}$  the set of intervals that are after  $x$ ,  $\mathcal{I}_{=x} = \{I \in \mathcal{I} : left(I) < x < right(I)\}$  the set of intervals that cross  $x$ , and  $\mathcal{I}_{\leq x} = \{I \in \mathcal{I} : right(I) \leq x\}$  the set of intervals that are before  $x$ .

Let us start with two lemmas that allow us to restructure optimal solutions by "flushing" intervals to the left.

**Lemma 11.** *Let  $S \subseteq \mathcal{I}$  be a solution, and  $s \in \mathbb{R}$  such that  $left(S) < s < right(S)$  and  $S \cap \mathcal{I}_{=s} = \emptyset$ . Let  $\tilde{I}$  be the leftmost interval of  $S \cap \mathcal{I}_{\geq s}$  and  $I^*$  be the leftmost interval of  $\mathcal{I}_{\geq s}$ . Then we can swap  $\tilde{I}$  and  $I^*$  to get a solution  $S' = (S \setminus \{\tilde{I}\}) \cup \{I^*\}$  such that  $cost(S') \leq cost(S)$ .*

*Proof.* Let us suppose that  $\tilde{I} \neq I^*$ , and let  $I \in S$  such that  $I \neq \tilde{I}, I^*$ . We will show that if  $I$  overlaps  $I^*$ , then it also overlaps  $\tilde{I}$ . Thus, suppose that  $I$  overlaps  $I^*$ . By definition of  $\tilde{I}$  and  $S$ , we have  $right(I^*) < right(\tilde{I}) < right(I)$ , and since  $I$  overlaps  $I^*$ , we have  $I \in \mathcal{I}_{=right(I^*)}$  and thus  $I$  also overlaps  $\tilde{I}$  (see Figure 3a).

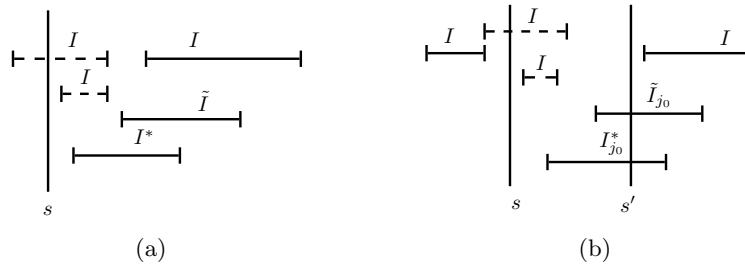


Fig. 3: Different positions of interval  $I$  in Lemma 11 (Figure (a)) and Lemma 13 (Figure (b)). Dashed intervals represent forbidden positions.

**Lemma 12.** *Let  $S \subseteq \mathcal{I}$  be a solution,  $I_{i_1} \in S$  and  $s \in \mathbb{R}$  such that:*

- (i)  $I_{i_1}$  is the leftmost interval of  $S \cap \mathcal{I}_{=s}$
- (ii)  $\exists \tilde{I} \in S \cap \mathcal{I}_{\geq s}$  such that  $\tilde{I}$  overlaps  $I_{i_1}$

Let  $I^*$  be the leftmost interval of  $\mathcal{I}_{\geq s}$ . Then, we can swap  $I^*$  and  $\tilde{I}$  to get a solution  $S' = (S \setminus \{\tilde{I}\}) \cup \{I^*\}$  such that  $\text{cost}(S') \leq \text{cost}(S)$ .

*Proof.* Let us suppose that  $\tilde{I} \neq I^*$ , otherwise the proof is obvious, and let  $I \in S$  such that  $I \neq \tilde{I}, I^*$ . We will show that if  $I$  overlaps  $I^*$ , then it also overlaps  $\tilde{I}$ . Thus, suppose that  $I$  overlaps  $I^*$ . If  $I \in \mathcal{I}_{=s}$ , then by definition of  $I_{i_1}$ , we must have  $\text{right}(I_{i_1}) < \text{right}(I)$ , and since  $s < \text{left}(\tilde{I}) < \text{right}(I_{i_1})$ ,  $I$  must overlap  $\tilde{I}$ . Otherwise if  $I \in \mathcal{I}_{\geq s}$ , as in the proof of Lemma 11, by definition of  $\tilde{I}$  we have  $\text{right}(I^*) < \text{right}(\tilde{I}) < \text{right}(I)$ , and since  $I$  overlaps  $I^*$ , we have  $I \in \mathcal{I}_{=\text{right}(I^*)}$  and thus  $I$  also overlaps  $\tilde{I}$ .

**Lemma 13.** Let  $S \subseteq \mathcal{I}$  be a solution and  $s, s' \in \mathbb{R}$  with  $s < s'$  and such that  $\forall I \in S$  we have  $\text{right}(I) \notin [s, s']$ . Let  $\tilde{X} = S \cap \mathcal{I}_{\geq s} \cap \mathcal{I}_{=s'}$  and  $X^*$  be the  $|\tilde{X}|$ -leftmost intervals of  $\mathcal{I}_{\geq s} \cap \mathcal{I}_{=s'}$ . Then we can swap  $\tilde{X}$  and  $X^*$  to get a solution  $S' = (S \setminus \tilde{X}) \cup X^*$  such that  $\text{cost}(S') \leq \text{cost}(S)$ .

*Proof.* We suppose that  $\tilde{X} \neq X^*$  and that both sets are non empty. Let  $\tilde{X} = \{\tilde{I}_1, \dots, \tilde{I}_{|\tilde{X}|}\}$ , and  $X^* = \{I_1^*, \dots, I_{|\tilde{X}|}^*\}$ . We suppose moreover that for all  $j \in \{2, \dots, |\tilde{X}|\}$  we have  $\text{right}(\tilde{I}_{j-1}) < \text{right}(\tilde{I}_j)$  and  $\text{right}(I_{j-1}^*) < \text{right}(I_j^*)$  (i.e.  $\tilde{X}$  and  $X^*$  are sorted by their right endpoints). Let  $j_0$  be the minimum index such that  $\tilde{I}_{j_0} \neq I_{j_0}^*$ , and let  $I \in S \setminus (\tilde{X} \cup X^*)$  (we thus have  $\text{right}(I_{j_0}^*) < \text{right}(\tilde{I}_{j_0})$ ). We will show that if  $I$  overlaps  $I_{j_0}^*$ , then  $I$  also overlaps  $\tilde{I}_{j_0}$ . To do so, suppose that  $I$  overlaps  $I_{j_0}^*$ , and let us distinguish between two cases (see Figure 3b). If  $s' < \text{right}(I)$ , then since  $\text{right}(I_{j_0}^*) < \text{right}(\tilde{I}_{j_0})$ , it is clear that  $I$  also overlaps  $\tilde{I}_{j_0}$ . Otherwise, if  $\text{right}(I) < s'$ , then by definition  $\text{right}(I) < s$ , and thus  $I$  cannot overlap  $I_{j_0}^*$ .

### 6.3 Algorithm

Recall that our objective is to prove that the decision problem "given an instance  $(\mathcal{I}, k)$  of SPARSEST  $k$ -SUBGRAPH, does  $\text{Opt}(\mathcal{I}, k) \leq \text{cost} ?$ ", is FPT parametrized by  $\text{cost}$ .

We construct in Algorithm 1 a dynamic programming algorithm that given any  $\text{next} \in \mathbb{R}$ ,  $t \leq k$  and  $C \leq \text{cost}$  returns a set  $S$  of  $t$  vertices in  $\mathcal{I}_{\geq \text{next}}$  of cost at most  $C$  if it was possible, and returns *NO* otherwise.

We define  $\Omega_{\text{next}}(C) \subseteq \mathcal{P}(\mathcal{I})$  (where  $\mathcal{P}(\mathcal{I})$  is the set of all subsets of  $\mathcal{I}$ ) such that for all  $T \in \Omega_{\text{next}}(C)$  we have:

- $G[T]$  is connected
- $\text{cost}(T) \leq C$
- $\text{left}(T) = \text{left}(\mathcal{I}_{\geq \text{next}})$

Roughly speaking,  $\Omega_{\text{next}}(C)$  is the set of all connected components of cost at most  $C$  that start immediately after  $\text{next}$ . Given  $\text{next}$  and  $t$ , the algorithm branches on a subset of  $\Omega_{\text{next}}(C)$  (namely  $\Gamma_{\text{next}}(C)$ ) to find what could be the next optimal connected component, and then invokes a recursive call.

We prove in Lemma 14 that each  $T \in \Omega_{\text{next}}(C)$  can be restructured into a "well-structured" component of smaller cost, and in Lemma 15 that the size of the set of all "well-structured" components ( $\Gamma_{\text{next}}(C)$ ) can be enumerated in *FPT* time.

**Algorithm 1**  $DP(next, t, C)$ 


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{F} for the sake of clarity we drop the classical operations related to the "marking
{t}able" that avoid multiple computations with same arguments
build  $\Gamma_{next}(C)$  (see Definition 1)
if  $\Gamma_{next}(C) = \emptyset$  then
    return NO
else if  $\exists T \in \Gamma_{next}(C)$  with  $|T| \geq t$  then
    return  $t$  vertices of  $T$ 
else
    return  $\arg \min_{C \in \Gamma_{next}(C)} [cost(T) + cost(DP(right(T), t - |T|, C - cost(T)))]$ 

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Let us now show how to restructure a connected component of a given solution. As one could expect, the idea is to apply the domination rules of Lemmas 11, 12 and 13 that consist in "flushing" the intervals to the left. Thus, for any connected component  $T$ , we define (recursively on  $s$ )  $restruct(s, T, i)$  that turns  $T \cap \mathcal{I}_{\geq s}$  (the part of  $T$  which is after  $s$ ) into a well structured solution (see Algorithm 2 and Definition 1). Notice that the parameter  $i$  and the  $y_i$  values will be used in Lemma 15 to show that the output of the algorithm can be encoded in an efficient way.

**Definition 1.** Given  $s$  and  $T \in \Omega_s(C)$ , we define:

- $WSS(T) = restruct(start(T), T, 0)$  the Well Structured Solution corresponding to  $T$ , where  $start(T)$  is defined as the point after the left endpoint of the leftmost interval of  $T$  (see Figure 4).
- $\Gamma_s(C) = \{WSS(T), T \in \Omega_s(C)\}$  the set of well structured connected component of cost at most  $C$  that starts just after  $s$ .

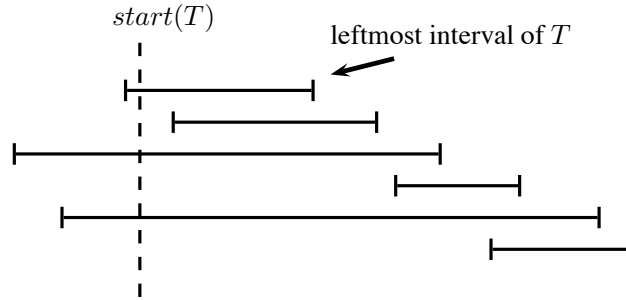


Fig. 4: Example of a connected component  $T$  and its corresponding  $start(T)$

*Remark 3.* Notice that at each step of the dynamic programming we branch on  $\Gamma_{next}(cost)$ , which is the set of all restructured connected component  $T$  such that  $left(C) = left(\mathcal{I}_{\geq next})$ . By Lemma 11, we can suppose that for all optimal solution  $S^*$ , we have  $left(S^* \cap \mathcal{I}_{\geq next}) = left(\mathcal{I}_{\geq next})$ . Roughly speaking, we can suppose that for all optimal solution, the connected component that starts after  $\mathcal{I}_{\geq next}$  contains the leftmost interval of  $\mathcal{I}_{\geq next}$ . As a consequence, the  $start(T)$  in Definition 1 forces  $I_{i_1}$  to be this leftmost interval.

**Algorithm 2**  $restruct(s, T, i)$ 


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if  $T \cap \mathcal{I}_{\geq s} \neq \emptyset$  then
   $I_{i_1} \leftarrow$  leftmost interval of  $\mathcal{I}_{=s} \cap T$ 
   $\{I_{i_1}$  is always defined, as in the first call  $s$  is set to  $start(T)$ 
  if  $\nexists I \in T \cap \mathcal{I}_{\geq s}$  which overlaps  $I_{i_1}$  then
     $y_i \leftarrow 0$ 
     $restruct(right(I_{i_1}), T, i + 1)$ 
  else
    {we restructure a first interval using Lemma 12}
     $\tilde{I} \leftarrow$  leftmost interval of  $T \cap \mathcal{I}_{\geq s}$  which overlaps  $I_{i_1}$ 
     $I^* \leftarrow$  leftmost interval of  $\mathcal{I}_{\geq s}$  which overlaps  $I_{i_1}$ 
     $T \leftarrow (T \setminus \{\tilde{I}\}) \cup \{I^*\}$ 
    {we restructure a set of intervals using Lemma 13}
     $s' \leftarrow \min(right(I^*), right(I_{i_1}))$ 
     $\tilde{X} \leftarrow T \cap \mathcal{I}_{\geq s} \cap \mathcal{I}_{=s'}$ 
     $X^* \leftarrow |\tilde{X}|$ -leftmost intervals of  $\mathcal{I}_{\geq s} \cap \mathcal{I}_{=s'}$ 
     $T \leftarrow (T \setminus \tilde{X}) \cup X^*$ 
     $y_i \leftarrow |X^*| + 1$ 
     $restruct(s', T, i + 1)$ 

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**Lemma 14.** For any  $s$  and any  $T \in \Omega_s(C)$  we have

- $|WSS(T)| = |T|$ , i.e. the restructured set has same size
- $right(WSS(T)) \leq right(T)$
- $cost(WSS(T)) \leq cost(T)$ .

*Proof.* The first item is clearly true as we only swap sets of intervals of same size. The second item is true as all swapping arguments shift intervals to the left. Let us now turn to the last item. Notice that in the two cases where  $restruct$  modifies  $T$ , the hypothesis of Lemmas 12 and 13 are verified. Thus, according to these Lemmas the cost of the solution cannot increase.

Lemma 14 confirms that the dynamic programming algorithm can only branch on  $\Gamma_s(C)$ , avoiding thus branching on  $\Omega_s(C)$ . It remains now to prove that the dynamic programming algorithm is FPT.

**Lemma 15.** For any  $s$ ,  $|\Gamma_s(C)| \leq (\sqrt{2C} + 2)^{C+1}$ .

*Proof.* Let  $T \in \Omega_s(C)$ , and  $WSS(T)$  the associated restructured solution. The key argument is to remark that  $WSS(T)$  is entirely determined by the  $y_i$  values defined in the  $restruct$  algorithm. Thus, to each restructured solution  $WSS(T)$  we associate the vector  $Y(WSS(T)) = (y_0, \dots, y_{l_{max}})$ . Then, the dynamic program will enumerate  $\Gamma_s(C)$  by enumerating the set  $Y = \{Y(WSS(T)), T \in \Omega_s(C)\}$  of all possible  $Y$  vectors.

Notice first that for any  $i$  we have  $y_i \leq \sqrt{2C} + 2$ . Indeed, in the two possible cases of the restructuration ( $s' = right(I^*)$  or  $s' = right(I_{i_1})$ ) the  $|X^*|$  intervals all overlap  $s'$ , corresponding to the right endpoint of another interval ( $I^*$  or  $I_{i_1}$ ). Thus, there is at least a clique of size  $y_i = |X^*| + 1$  in the solution, whose cost is lower than  $C$ .

It remains now to bound  $l_{max}$ , the length of the  $Y$  vector.

To do that, we show that for any step  $i \in \{0, \dots, l_{max} - 1\}$  and corresponding  $s$ , we can find  $I \in \mathcal{I}_{=s}$  and  $I' \in \mathcal{I}_{>s}$  such that  $I$  and  $I'$  overlaps, and such that in the next recursive call (with parameter  $s'$ ), either  $I$  or  $I'$  belongs to  $\mathcal{I}_{\leq s'}$ , avoiding multiple counts of same pairs, and implying that  $C \geq l_{max} - 1$ . Let  $i \in \{0, \dots, l_{max} - 1\}$ . If  $y_i \neq 0$ , then by definition of  $I^*$ ,  $I_{i_1}$  and  $I^*$  are overlapping. Then, since the next recursive call has parameter  $s' = \min\{right(I^*), right(I_{i_1})\}$ , either  $I^*$  or  $I_{i_1}$  belongs to  $\mathcal{I}_{\leq s'}$ . If  $y_i = 0$ , then  $s' = right(I_{i_1})$ , and as  $i \neq l_{max}$ , we know that there exists  $I_{i_2} \in \mathcal{I}_{=s'}$  implying that  $I_{i_2}$  overlaps  $I_{i_1}$ . Finally, it is clear that  $I_{i_1} \in \mathcal{I}_{\leq s'}$ .

**Theorem 4.** SPARSEST  $k$ -SUBGRAPH can be solved in  $O(n^2 \cdot k^3 \cdot cost \cdot (\sqrt{2cost} + 2)^{cost+1})$ .

*Proof.* The dynamic programming algorithm has at most  $n \cdot k \cdot cost$  different inputs. Given fixed parameters, it runs in  $O(|\Gamma_s(cost)| \cdot k^2 n)$ . Indeed, given a  $Y$  vector, the corresponding connected component can be built in  $O(l_{max} n) \subseteq O(n \cdot cost) \subseteq O(k^2 n)$  as for any  $i \leq l_{max}$  it takes  $O(n)$  to find the corresponding  $y_i$  intervals.

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## A Formal Definitions for Kernel Lower Bounds

In order to establish kernel lower bounds, we use the technique of cross-composition of [2]:

**Definition 2 (Polynomial equivalence relation [2]).** An equivalence relation  $\mathcal{R}$  on  $\Sigma^*$  is called a polynomial equivalence relation if the two following conditions hold:

- There is an algorithm that given two strings  $x, y \in \Sigma^*$ , decides whether  $x$  and  $y$  belong to the same equivalence class in  $(|x| + |y|)^{O(1)}$  time.
- For any finite set  $S \subseteq \Sigma^*$ , the equivalence relation  $\mathcal{R}$  partitions the elements of  $S$  into at most  $(\max_{x \in S} |x|)^{O(1)}$  classes.

**Definition 3 (OR-cross-composition [2]).** Let  $L \subseteq \Sigma^*$  be a set and let  $Q \subseteq \Sigma^* \times \mathbb{N}$  be a parameterized problem. We say that  $L$  OR-cross-composes into  $Q$  if there is a polynomial equivalence relation  $\mathcal{R}$  and an algorithm which, given  $t$  strings belonging to the same equivalence class of  $\mathcal{R}$ , computes an instance  $(x^*, k^*) \in \Sigma^* \times \mathbb{N}$  in time polynomial in  $\sum_{i=1}^t |x_i|$  such that:

- $(x^*, k^*) \in Q \Leftrightarrow x_i \in L$  for some  $1 \leq i \leq t$
- $k^*$  is bounded by a polynomial in  $\max_{i=1}^t |x_i| + \log t$

**Theorem 5 ([2]).** If some set  $L \subseteq \Sigma^*$  is  $\mathcal{NP}$ -hard and  $L$  OR-cross-composes into the parameterized problem  $Q$ , then there is no polynomial kernel for  $Q$  unless  $\mathcal{NP} \subseteq \text{coNP}/\text{poly}$ .