# DEPARTMENT OF INFORMATICS

TECHNICAL UNIVERSITY MUNICH

Master Thesis

# On the Parameterized Complexity of Semitotal Domination on Graph Classes

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# Über die Parametrisierte Komplexität des Problems der halbtotalen stabilen Menge auf Graphklassen

| | |
|---|---|
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| Submission Date: | December 19, 2022 |

I confirm that this master thesis is my own work and I have documented all sources and material used.

*København S*                                                    Lukas Retschmeier
December 19, 2022

# Acknowledgments

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# ABSTRACT

### Abstract

For a graph $G = (V, E)$, a set $D$ is called a *semitotal dominating set*, if $D$ is a dominating set and every vertex $v \in D$ is within distance two to another witness $v' \in D$. The MINIMUM SEMITOTAL DOMINATING SET problem is to find a semitotal dominating set of minimum cardinality. The semitotal domination number $\gamma_{t2}(G)$ is the minimum cardinality of a semitotal dominating set and is squeezed between the domination number $\gamma(G)$ and the total domination number $\gamma_t(G)$. Given a graph $G = (V, E)$ and a positive integer $k$, the SEMITOTAL DOMINATION DECISION problem asks if $G$ has a semitotal dominating set of size at most $k$.

After the problem was introduced by Goddard, Henning and McPillan in [15], **NP**-completeness was shown for general graphs [21], *split graphs* [21], *planar graphs* [21], *chordal bipartite graphs* [21], *circle graphs* [24] and *subcubic line graphs of bipartite graphs* [13]. On the other side, there exist polynomial-time algorithms for *AT-free graphs* [24], *graphs of bounded mim-width* [13], *graphs of bounded clique-width* [5], and *interval graphs* [21].

In this thesis, we start the systematic look through the lens of *parameterized complexity* by showing that SEMITOTAL DOMINATING SET is $\omega[2]$-hard for bipartite graphs and split graphs. By applying the techniques proposed in [1] and [14] for DOMINATING SET and TOTAL DOMINATING SET, we are going to construct a $359k$ kernel for SEMITOTAL DOMINATING SET in planar graphs. This result further complements known linear kernels for other domination problems like PLANAR CONNECTED DOMINATING SET, PLANAR RED-BLUE DOMINATING SET, PLANAR EFFICIENT DOMINATING SET, PLANAR EDGE DOMINATING SET, INDEPENDENT DOMINATING SET and PLANAR DIRECTED DOMINATING SET.

**Keywords:** Domination; Semitotal Domination; parameterized Complexity; Planar Graphs; Linear Kernel
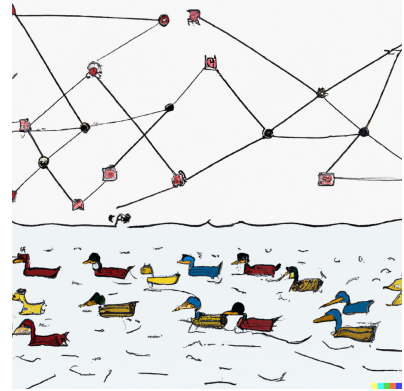
# ZUSAMMENFASSUNG

### Abstract

Hier kommt noch ein weiterer Abstract rein.

**Schlagworte**: Stabile Menge; Halbtotale Stabile Menge; Parametrisierte Komplexität; Plättbare Graphen; Linearer Problemkern

# CHAPTER 1

## TERMINOLOGY AND PRELIMINARIES



> *"All we have to decide is what to do with the time that is given to us."*
>
> J. R. R. Tolkien, *Gandalf* in *Lord of the Rings*

In this chapter, we will introduce the core definitions used throughout this thesis. Most of the definitions of graph theory are taken from [7]. For definitions in the area of *parameterized complexity*, the book written by Cygan et al. [6] gives an excellent introduction. For standard mathematical notation, the reader is referred to any introductory textbook into discrete mathematics (e.g. [27]).

## 1.1 Graph Theory

If not explicitly stated otherwise, the following definitions are taken from the book *Graph Theory* written by Reinhard Diestel [8].

### 1.1.1 Basic Terminology

**Definition 1.1.1** (Graph). *A simple graph is a pair $G = (V, E)$ of two sets where $V$ denotes the vertices and $E \subseteq V \times V$ the edges of the graph. A vertex $v \in V$ is incident with an edge $e \in E$ if $v \in e$. Two vertices $x, y$ are adjacent, or neighbours, if $\{x, y\} \in E$. By this definition, graph loops and multiple edges are excluded.*

*A multigraph is a pair $(V, E)$ of disjoint sets together with a map $E \to V \cup [V]^2$ assigning to every edge either one or two vertices, its ends. Multigraphs can have loops and multiple edges. We usually denote the vertex set by $V(G)$ and its edge set by $E(G)$.*

Unless stated otherwise, we usually consider only *simple graphs*, but the notion of *multigraphs* gets important when we later talk about the *underlying multigraph* of a *D-region decomposition*.

**Definition 1.1.2** (Subgraph and Induced Subgraph). *Let $G = (V, E)$ and $G' = (V', E')$ be two graphs. If $V' \subseteq V$ and $E' \subseteq E$ then $G'$ is a subgraph of $G$. If $G$ is a subgraph of $G'$ and $G'$ contains all the edges to $G$ with both endpoints in $V(G')$, then $G'$ is an induced subgraph of $G$ and we write $G' = G[V(G')]$.*

**Definition 1.1.3** (Degrees). *Let $G = (V, E)$ be a graph. The degree $d_G(v)$ (shortly $d(v)$ if $G$ is clear from the context) of a vertex $v \in V$ is the number of neighbors of v. We call a vertex of degree $0$ as isolated and one of degree $1$ as a pendant. If all the vertices of $G$ have the same degree $k$, then $g$ is $k$-regular.*

**Definition 1.1.4** (Closed and Open Neighborhoods [3]). *Let $G = (V, E)$ be a (non-empty) graph. The set of all neighbors of $v$ is the open neighborhood of $v$ and denoted by $N(v)$; the set $N[v] = N(v) \cup \{v\}$ is the closed neighborhood $f$ $v$ in $G$. When $G$ needs to be made explicit, those open and closed neighborhoods are denoted by $N_G(v)$ and $N_G[v]$.*

**Definition 1.1.5** (isomorphic Graphs). *Let $G = (V, E)$ and $G' = (V', E')$ be two graphs. We call $G$ and $G'$ isomorphic, if there exists a bijection $\phi : V \to V'$ with $\{x, y\} \in E \Leftrightarrow \phi(x)\phi(y) \in E'$ for all $x, y \in V$. Such a map $\phi$ is called isomorphism.*
*If a graph $G$ is isomorphic to another graph $h$, we denote $G \simeq H$.*

**Definition 1.1.6** (Paths and Cycles). *A path is a non-empty graph $P = (V, E)$ of the form $V = \bigcup_{i \in [k]} \{x_i\}$ and $E = \bigcup_{i \in [k-1]} \{x_i x_{i+1}\}$ where the $x_i$ are distinct. The vertices $x_0$ and $x_k$ are linked by $P$ and are called the ends of $P$. The length of a path is its number of edges and the path on $n$ vertices is denoted by $P_n$. We refer to a path $P$ by a natural sequence of its vertices: $P = x_0 x_1 ... x_k$. Such a path $P$ is a path between $x_0$ and $x_k$, or a $x_0, x_k$-path. If $P = x_0 ... x_k$ is a path and $k \geq 2$, the graph with vertex set $V(P)$ and edge set $E(P) \cup \{x_k x_0\}$ is a cycle. The cycle on $n$ vertices is denoted as $C_n$. The distance $d_G(v, w)$ from a vertex $v$ to a vertex $w$ in a graph $g$ is the length of the shortest path between $v$ and $w$. If $v$ and $w$ are not linked by any path in $G$, we set $d_G(v, w) = \infty$. Again, if $G$ is clear from the context, we omit the subscripted $G$ and just write $d(v, w)$ instead.*

### 1.1.2 Graph Classes

A *graph class* is a set of graphs $\mathfrak{G}$ that is closed under isomorphism that is if $G \in \mathfrak{G}$ and a $H \simeq G$ then $H \in \mathfrak{G}$ as well.

**Definition 1.1.7** (Graph Parameters). *Let $G = (V, E)$ be a graph. An independent set of G is a set of pairwise non-adjacent vertices. A clique of G is a set of pairwise adjacent vertices. A vertex cover of G is a subset of vertices containing at least one endpoint of every edge. A dominating set is a subset D of vertices such that all vertices not contained in are adjacent to some vertex in D.*

**Graph Class 1** (r-partite). *Let $r \geq 2$ be an integer. A Graph $G = (V, E)$ is called r-partite if V admits a partition into r classes such that every edge has its ends in different classes: Vertices in the same partition class must not be adjacent. A 2-partite graph is called bipartite.*

*An r-partite graph in which every two vertices from different partition classes are adjacent is called complete. For the complete bipartite graph on bipartitions $X \uplus Y$ of size m and n, we shortly write $K_{m,n}$.*

**Graph Class 2** (Complete). *If all vertices of a graph $G = (V, E)$ are pairwise adjacent, we say that G is complete. A complete graph on n vertices is a $K_n$. A $K_3$ is called a triangle.*

**Graph Class 3** (Chordal). *For a graph $G = (V, E)$, an edge that joins two vertices of a cycle, but is not itself an edge of the cycle is a chord of that cycle.*

*Furthermore, we say G is chordal (or triangulated) if each of its cycles of length at least four has a chord. In other words, it contains no induced cycle other than triangles.*

**Graph Class 4** (Split). *A split graph is a graph $G = (V, E)$ whose vertices can be partitioned into a clique and an independent set.*

**Graph Class 5** (Planar). *A plane graph is a pair $(V, E)$ of finite sets with the following properties:*

- *$V \subseteq \mathbb{R}^2$ (Vertices),*
- *Every edge is an arc between two vertices,*
- *different edges have different sets of endpoints, and*
- *The interior of an edge contains no vertex and no point of any other edge*

*An embedding in the plane, or planar embedding, of an (abstract) graph G is an isomorphism between G and a plane graph H. A plane graph can be seen as a concrete **embedding** of the planar graph into the "plane" $\mathbb{R}^2$.*

## 1.2 Computational Complexity Theory

Computational complexity investigates the question of how many computational resources are required to solve a specific problem. We are about to introduce two of the most important classes of problems in classical complexity theory:

> **The Class P [2]**
>
> If we denote **DTIME** as the set of decision problems that are solvable in $\mathcal{O}(n^k)$ time by a deterministic Turing Machine, we can define the class **P** as:
>
> $$P := \bigcup_{k \in \mathbb{N}} (\text{DTIME}(n^k))$$

> **The Class NP [2]**
>
> A language $L \subseteq \{0,1\}^*$ is in **P** if there exists a polynomial $p : \mathbb{N} \to \mathbb{N}$ and a polynomial-time Turing Machine $M$ such that for every $x \in \{0,1\}^*$,
>
> $$x \in L \Leftrightarrow \exists u \in \{0,1\}^{p(|x|)} s.t.\ M(x,u) = 1$$
>
> If $x \in L$ and $u \in \{0,1\}^{p(|x|)}$ satisfy $M(x,u) = 1$, then we call $u$ a *certificate* for $x$.

**P** denotes the class of all problems that are *efficiently solvable* whereas **NP** contains all problems whose solution can efficiently be verified. Note that **P** $\subseteq$ **NP**, but the opposite is unknown.

### 1.2.1 NP-Completeness

A major discovery in the early 1970s was the fact that some problems in **NP** are *at least as hard as* as any other problem in **NP** by reducing them among each other. This spans a whole "web of reductions" [2] and gives strong evidence that none of these problems can be solved efficiently. The first results in this new field had been published independently by Cook [4] and Levin [25] after Karp [23] had introduced this notion of problem reductions. The Cook-Levin-Theorem [4] proved that the BOOLEAN SATISFIABILITY (SAT) problem is **NP-COMPLETE** any problem in $NP$ can be reduced to SAT.

as hard as SAT, because a fast algorithm for $P$ would immediately give a fast algorithm for SAT as well. one single algorithm for any of these problems would be enough to efficiently solve all of them. For a comprehensive introduction to classical complexity theory, the reader is referred to [2].

**Definition 1.2.1** (Reductions, **NP**-hardness and **NP-COMPLETENESS** [2])**.** *We say that a language $A \subseteq \{0,1\}^*$ is polynomial-time Karp reducible to a language $B \subseteq \{0,1\}^*$ (denote $A \leq_p B$) if there is a poly-time computable function $f : \{0,1\}^* \to \{0,1\}^*$ such that for every $x \in \{0,1\}^*$, $x \in A$ if and only if $f(x) \in B$.*
*We say that a problem $B$ is **NP-HARD** if $A \leq_p B$ for every $A \in$ **NP** and $B$ is **NP-COMPLETE** if additionally $B \in NP$ holds.*

There are thousands of **NP**-Complete problems we do not expect to be solvable in polynomial time. The famous question of whether **P** = **NP** or not is still one of the biggest open questions in mathematics bountied with one million dollars by the *Clay Mathematical Institute* [12]. Most of the domination problems like Dominating Set, Semitotal Dominating Set, Total Dominating Set are **NP**-Complete.

**Coping with NP-Completeness**   Even though we do not expect **NP**-Complete problems to have a polynomial-time algorithm, there are some strategies to cope with them. We can either give up the exactness of a solution to possibly find fast *approximation algorithms* or abandon the search for a polynomial-time algorithm in favor of finding good *Exact Exponential (EEA) Algorithms* instead.

A third technique is using additional structural parameters of a specific problem instance and therefore **restricting the input to special cases**. This idea lead to the development of *Parameterized complexity*.

### 1.2.2 Definitions in Parameterized Complexity

Introduced by Downey and Fellows [9], parameterized complexity extends the classical theory with a framework that allows a more finely-grained analysis of computationally hard problems. The idea is to measure a problem in terms of input size and an additional (structural) parameter $k$.

We like to find an algorithm that is only exponential in a function $f(k)$, but polynomial in the instance size. $k$ denotes how difficult the problem is:

If $k$ is small then the problem can still be considered tractable although the underlying **NP**-hard problem counts as intractable in general. Therefore $k$ can be seen as a measure of the difficulty of a given instance. If not marked otherwise, all definitions are taken from [6].

**Definition 1.2.2** (Parameterized Problem). *A parameterized problem is a $L \subseteq \Sigma^* \times \mathbb{N}$ ($\Sigma$ finite fixed alphabet) for an instance $(x, k) \in \Sigma^* \times \mathbb{N}$, where $k$ is called the parameter.*

*The size of an instance of an instance $(x, k)$ of a parameterized problem is $|(x, k)| = |x| + k$ where the parameter $k$ is encoded in unary by convention.*

### 1.2.3 Fixed-Parameter Tractability

We say that a problem is *fixed-parameter tractable (fpt)* if problem instances of size $n$ can be solved in $f(k)n^{\mathcal{O}(1)}$ time for some function $f$ independent of $n$. Like the class **P** can be seen as a notion of *tractability* in classical complexity theory, there is an equivalent in parameterized complexity, which we denote as Fixed-Parameter Tractable (**FPT**) and which we can define the following way:

> **The Class FPT**
>
> A parameterized problem $L \subseteq \Sigma^* \times \mathbb{N}$ is called *fixed-parameter tractable* if there exists an algorithm A (called a *fixed-parameter algorithm*), a computable function $f : \mathbb{N} \to \mathbb{N}$ and a constant c such that, given $(x, k) \in \Sigma^* \times \mathbb{N}$, the algorithm $\mathcal{A}$ correctly decides whether $(x, k) \in L$ in time bounded by $f(k) \cdot |(x, k)|^c$. The complexity class containing all fixed-parameter tractable problems is called **FPT**.

### 1.2.4 Kernelization

A kernelization algorithm is a natural and intuitive way to approach problems and can be seen as a preprocessing procedure that simplifies parts of an instance already before the actual solving algorithm is run. A visualization of this idea can be seen in Figure 1.2. One can introduce *reduction rules* that iteratively reduce the instance until we are left with a small kernel.

**Definition 1.2.3** (Kernelization and Reduction Rules). *A kernelization algorithm or kernel is an algorithm $\mathfrak{A}$ for a parameterized problem Q that given an instance $(I, k)$ of Q runs in polynomial time and returns an equivalent instance $(I', k')$ of Q. Moreover, we require that $size_{\mathfrak{A}}(k) \leq g(k)$ for some computable function $g : \mathbb{N} \to \mathbb{N}$.*
   *A <u>reduction rule</u> is a function $\phi : \Sigma^* \times \mathbb{N} \to \Sigma^* \times \mathbb{N}$ that maps an instance $(x, k)$ to an equivalent instance $(x', k')$ such that $\phi$ is computable in time polynomial in $|x|$ and k.*
   *A reduction rule is <u>sound</u> (or <u>safe</u>) if $(I, k) \in Q \Leftrightarrow (I', k') \in Q$.*

We can give a precise definition of the size of the kernel, after a preprocessing algorithm $\mathfrak{A}$ has been executed. $size_{\mathfrak{A}}$ denotes the largest size of any instance $I$ after $\mathfrak{A}$ has been applied. We consider the size to be infinite if it cannot be bounded by a function in $k$.

**Definition 1.2.4** (Output size of a Preprocessing Algorithm). *The output size of a preprocessing algorithms $\mathfrak{A}$ is defined as*

$$\text{size}_{\mathfrak{A}}(k) = \sup\{|I'| + k' : (I', k') = \mathfrak{A}(I, k), I \in \Sigma^*\}$$

If we bound $\text{size}_{\mathfrak{A}}$ by a polynomial in $k$, we say that the problem admits a **polynomial kernel**. Analogous, if the size after the reduction is only linear $k$, we refer to it as a **linear** kernel.

The following Lemma 1.2.1 shows the relation between the complexity class **FPT** and a kernelization algorithm. If we find a kernelization algorithm $\mathfrak{A}$ for a (decidable) problem $P$, we immediately obtain an fpt algorithm by first running the $\mathfrak{A}$ on an instance $I$ of $P$ in polynomial time. Assuming that $P$ can be solved by an algorithm $\mathfrak{M}$ running in time $g(n)$ we can use the fact that the kernel is bounded by a function $f(k)$ and apply
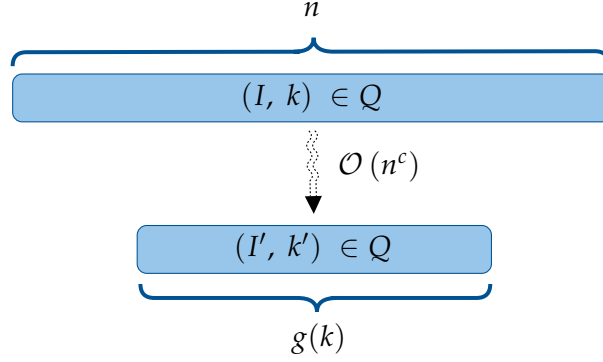
$$n$$



Figure 1.2: *Kernelization: Reducing an instance $(I, k)$ of size n to a smaller instance $(I', k')$ in polynomial time. The size of the kernel is a function $g(k)$ only dependent on k.*

$\mathfrak{M}$ on the kernel resulting in a total running time of the order $\mathcal{O}(g(f(k)) \cdot \text{poly}(n))$ which is fpt. Surprisingly, also the converse is true:

**Lemma 1.2.1.** *If a parameterized problem Q is* **FPT** *if and only if it admits a kernelization algorithm.*

In **??** we will use this and by explicitly constructing a kernel for PLANAR SEMITOTAL DOMINATING SET, we show membership of the problem in **FPT**.

### 1.2.5 Reductions and Parameterized Intractability

It is natural to ask whether all (hard) problems are also fixed-parameter tractable. The answer is no and parameterized complexity has another tool in its toolbox that can be used to show that a problem is unlikely to be in **FPT**. The idea is to transfer the concepts of **NP**-hardness from Section 1.2.1 and reductions from the classical setting to the parameterized world. This raises the need for a new type of reduction that ensures that a reduced instance $(I', k')$ is not only created in fpt time, but the new parameter $k'$ depends only on the size of the parameter in the original instance.

There exists a whole hierarchy of classes **FPT** $\subseteq$ **W[1]** $\subseteq$ **W[2]** $\subseteq$ ... $\subseteq$ **W[t]** $\subseteq$ ..., which is known as the **W**-hierarchy. It is strongly believed that **FPT** $\subsetneq$ **W[t]** and therefore, we do not expect the existence of an algorithm solving any **W[t]**-hard problem in fpt time.

**Definition 1.2.5** (Parameterized Reduction). *Let $A, B \subseteq \Sigma^* \times \mathbb{N}$ two parameterized problems. A parameter preserving reduction from A to B is an algorithm that, given an instance $(x, k)$ of A, outputs an instance $(x', k')$ of B such that:*

- *$(x, k)$ is a **yes instance** of A **iff** $(x', k')$ is a **yes instance** of B,*

- *$k' \leq g(k)$ for some computable function g, and*

- *runs in fpt-time $f(k) \cdot |x|^{\mathcal{O}(1)}$ for some computable function f.*

As shown in Lemmas 1.2.2 and 1.2.3 [6] this definition ensures that reductions are transitive and closed under fpt reductions.

**Lemma 1.2.2** (Closed under fpt-reductions). *If there is a parameterized reduction from A to B and B ∈ **FPT**, then A ∈ **FPT**, too.*

**Lemma 1.2.3** (Transitivity). *If there are parameterized reductions from A to B and from B to C, then there is a parameterized reduction from A to C.*

If there exists a parameterized reduction transforming a **W**[**t**]-hard problem $A$ to another problem $B$, then $B$ is **W**[**t**]-hard as well. We can define the classes **W**[**1**] and **W**[**2**], by giving two problems that are complete for these classes.

---

**The Classes W[1] [10] and W[2] [6]**

INDEPENDENT SET is **W**[**1**]-complete.

DOMINATING SET is **W**[**2**]-complete.

A problem $P$ is in the class **W**[**1**] (resp. **W**[**2**]) if there is a parameterized reduction from $P$ to INDEPENDENT SET (DOMINATING SET).

---

We have omitted a more precise definition via the WEIGHTED BOOLEAN SATISFIABILIY problem as it is not important for our work. We refer the interested reader to [6, 11] for more details.

# ON PARAMETERIZED SEMITOTAL DOMINATION



*TODO: Find a nice quote for here!*

chatGPT, 2022

In connection with various chessboard problems, the concept of domination can be traced back to the mid-1800s. For example, De Jaenosch attempted in 1862 to solve the minimum number of queens required to fully cover an $n \times n$-chessboard [22]. Because of the immense amount of publications related to domination, Haynes, Hedetniemi, and Slater started a comprehensive survey of the literature in 1998 [16, 17]. 20 years later, by a series of three more books, Haynes, Henning and Hedetniemi complemented the survey with the latest developments [18, 19, 20].

We are now introducing the problems of DOMINATING SET, SEMITOTAL DOMINATING SET and TOTAL DOMINATING SET and dedicate the rest of the chapter to giving a current status about the complexity status of various graph classes.

## 2.1 Definition of Domination Problems

We are now going to define the three most important domination problems for this work: DOMINATING SET, SEMITOTAL DOMINATING SET and TOTAL DOMINATING SET. For a specific dominating set $D$, we say that a vertex $d \in D$ is a *dominating vertex* or *dominator*. We say that *d dominates* all of its neighbors.

The DOMINATING SET problem asks for a subset $D$ of vertices whose set of neighbors covers all the other vertices. In other words: every vertex $v \notin D$ needs to have at least one neighbor in $D$.

---

**DOMINATING SET [6, p. 586]**

| | |
|---|---|
| **Input** | Graph $G = (V, E)$, $k \in \mathbb{N}$ |
| **Question** | Is there a set $X \subseteq V$ of size at most $k$ such that $N[X] = V$? |

---

The TOTAL DOMINATING SET problem adds one additional constraint: Every vertex $v \in D$ in the dominating set must also be dominated by some vertex $v' \in D$ which we call the *witness* of $v$.

---

**TOTAL DOMINATING SET [6, p. 596]**

| | |
|---|---|
| **Input** | Graph $G = (V, E)$, $k \in \mathbb{N}$ |
| **Question** | Does there exists a set $X \subseteq V$ with $|X| \leq k$ vertices such that for every $u \in V(G)$ there exists $v \in X$ with $\{u, v\} \in E$ |

---

Finally, SEMITOTAL DOMINATION was introduced by Goddard, Henning and McPillan [15] as a relaxation of TOTAL DOMINATION. Assume that we have an arbitrary dominating set $D$ for some (connected) graph $G = (V, E)$ that has size at most two.

It is easy to observe that for each $v \in D$ there must be at least one other dominating vertex $v' \in D$ that is at most three steps away because otherwise, this would not be a dominating set. By definition, there is always another neighboring dominating vertex for a total dominating set $D$. Therefore, it is natural to ask what happens, if we restrict the distance to at most two, which leads us straight to the idea of SEMITOTAL DOMINATION.

to still enforce another vertex from $D$ to be nearby, but allow to have intermediate In this context, we say that $v$ witnesses $v'$ if $v, v' \in D$ and $d(v, v') \leq 2$ for a semitotal dominating set $D$.

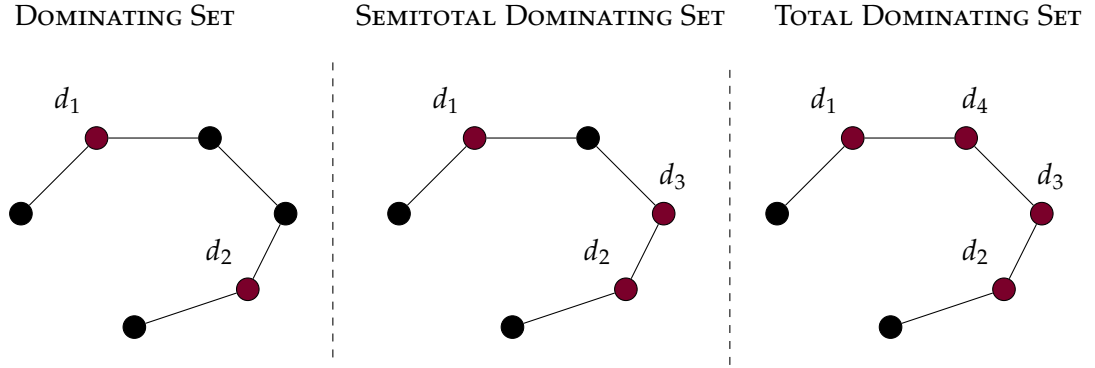DOMINATING SET     SEMITOTAL DOMINATING SET     TOTAL DOMINATING SET



Figure 2.2: *An example for a dominating set, semitotal dominating set and a total dominating set, where $\gamma(G) < \gamma_{2t}(G) < \gamma_t(G)$ are strict. In the first case, only two vertices suffice to dominate all others. In the second one, we need a witness between $d_1$ and $d_2$ that is at most distance two. In the last case, $d_1$ and $d_2$ both need a neighbor in the total dominating set.*

---

**SEMITOTAL DOMINATING SET [15]**

| | |
|---|---|
| **Input** | Graph $G = (V, E)$, $k \in \mathbb{N}$ |
| **Question** | Is there a subset $X \subseteq V$ with $|X| \leq k$ such that $N[X] = V$ and for all $d_1 \in X$ there exists another $d_2 \in X$ such that $d(d_1, d_2) \leq 2$? |

---

**Definition 2.1.1** (Domination Parameters)**.** *The <u>domination number</u> in a graph G is the minimum cardinality of a dominating set of G, denoted as $\gamma(G)$. The <u>total domination number</u> is the minimum cardinality of a total dominating set (tds) of G, denoted by $\gamma_t(G)$. The <u>semitotal domination number</u> is the minimum cardinality of a semitotal dominating set (sds) of G, denoted by $\gamma_t(G)$.*

Since every total dominating set is also a semitotal dominating set and every semitotal dominating set is also a dominating set , we have the following fact first observed by Goddard and Henning [15].

**Fact 2.1.1.** *For every graph G with no isolated vertex, $\gamma(G) \leq \gamma_{t2}(G) \leq \gamma_t(G)$*

We can see that the semitotal domination number $\gamma_{t2}$ is squeezed between the *domination* number and the *total domination* number. It turns out that for some graphs, all of these inequalities can be strict. See Figure 2.2 for an example, where $\gamma(G) < \gamma_{t2} < \gamma_t(G)$.

### 2.1.1 Preliminaries

* Witness * u pendant ofrom a vertex c if $N(u) = \{w\}$ * domination

Let $D$ be a dominating set of G and $w \in V(G) \setminus D$. For any neighbor $v \in D \cap N(w)$, we say that $d_1$ *dominates w* For two dominating vertices $d_1, d_2 in D$. If

Definition, dominating number

## 2.2 Complexity Status of Semitotal Dominating Set

## 2.3 *w[i]*-Intractibility

Now some w[i] hard classes.

### 2.3.1 Warm-Up: *W[2]*-hard on General Graphs

As any bipartite graphswith bipartition can be split further into *r*-partite graphsthis results also implies the $w[1]$-hardness of *r*-partite graphs

### 2.3.2 *W[2]*-hard on Bipartite Graphs

We are showing that SEMITOTAL DOMINATING SET is $\omega[2]$-hard on bipartite graphs by a parameterized reduction from DOMINATING SET on bipartite graphs which is known to be $\omega 2$-hard ([26, Theorem 1]).
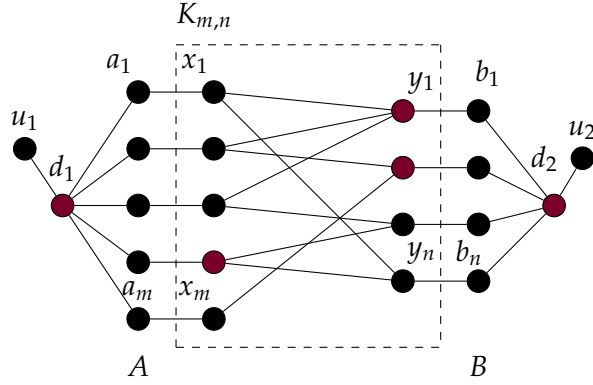


Figure 2.3: *Constructing a bipartite G' from the bipartite graph $K_{m,n}$ by duplicating all vertices and adding exactly two forced witnesses.*

**Theorem 1.** SEMITOTAL DOMINATING SET *is* $\omega[2]$ *hard for bipartite Graphs*

*Proof.* Given a bipartite Graph $G = (\{X \cup Y\}, E)$, we construct a bipartite Graph $G' = (\{X' \cup Y'\}, E')$ in the following way:

1. For each vertex $x_i \in X$, we add a new vertex $a_i \in A$ and an edge $\{x_i, a_i\}$ in between.

2. For each vertex $y_j \in Y$, we add a new vertex $b_j \in B$ and an edge $\{y_j, b_j\}$ in between.

3. We add fource vertices with with edges $\{u_1, d_1\}$ and $\{u_2, d_2\}$, and connect them with all $\{d_1, a_i\}$ and $\{d_2, b_j\}$ ($i \in [m]$ and $j \in [n]$) respectively.

**Observation:** The constructed $G'$ is bipartite because $A$ and $B$ form an independent set on $G'$ that can be cross-wise attached to one of the previous vertex sets. Setting $X' = X \cup \{u_2, d_1\} \cup B$ and $Y' = X \cup \{u_1, d_2\} \cup A$ form the partitions of the new bipartite $G'$.

**Corollary 2.3.1.** *G has a dominating set of size k iff G has a semitotal dominating set of size* $k' = k + 2$

$\Rightarrow$ Assume a ds in $G$ of size $k$. We know that $D' = D \cup \{d_1, d_2\}$ is an sds in $G'$ of size $k' = k + 2$, because $d_1$ dominates $u_1$ and all $a_i \in A$; $d_2$ dominates $u_2$ and all

$b_i \in B$. The rest is dominated by the same vertices as they were in $G$, but now all of them have either $d_1$ or $d_2$ as a witness. More formally, we have by construction of $G'$ that $\forall v \in (D \cap X) : d(v, d_1) = 2$ and $\forall v \in (D \cap Y) : d(v, d_2) = 2$.

$\Leftarrow$ On the contrary, assume any sds $D'$ in $G'$ with size $k'$. Without loss of generality, we can assume that $u_1, u_2 \notin D'$, because choosing $d_1$ and $d_2$ instead is always as least as good and does not violate any witnesses. Therefore, the construction forces $d_1, d_2 \in D'$.

All $a_i \in A$ can only be useful to dominate their partnering $x_i$ ($b_i \in B$ for $y_i$), because $d_1, d_2 \in D$ is the only second neighbor they have. If $a_i, b_i \in D'$ we replace it with $x_i$ and $y_i$ preserving the size $D$. As $d_1$ and $d_2$ suffice to provide a witness for every vertex in the graph and do not lose any other witnesses, this operation is sound.

In the end, $D = D' \setminus \{d_1, d_2\}$ gives us a ds in $G$ with size $k = k' - 2$

As $G'$ can be constructed in linear time and the parameter $k$ is only blown up by a constant, this reduction is an FPT reduction. Because DOMINATING SET is already $w[2]$-hard on bipartite graphs ([]), we imply that SEMITOTAL DOMINATING SET is $w[2]$-hard as well. $\square$

### 2.3.3 *W[2]-hard on Split Graphs*

TODO Getting started with that.

### 2.3.4 *W[2]-hard on Chordal Graphs*

Although the previous result implies $w[2]$-hardness for chordal graphs, we found another reduction from DOMINATING SET on chordal graphs.

We will introduce the notion of an elimination ordering.

**Definition 2.3.1 ([Rose1960]).** *In a graph $G = (V, E)$ with $n$ vertices, a vertex is called* ***simplicial*** *if and only if the subgraph of $G$ induced by the vertex set $\{v\} \cup N(v)$ is a complete graph.*

*G is said to have a* ***perfect eliminiation ordering*** *if and only if there is an ordering $(v_1, ...v_n)$ of the vertices, such that each $v_i$ is simplicial in the subgraph induced by the vertices $v_1, ..., v_i\{\}$*

The following lemma shows that

**Lemma 2.3.1 ([Rose1960]).** *A graph $G = (V, E)$ is chordal if and only if $G$ has a perfect elimination ordering.*

**Theorem 2.** SEMITOTAL DOMINATING SET *restricted to chordal graphs is $\omega[2]$-hard.*
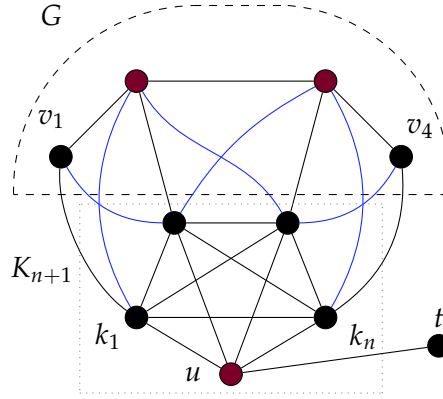
Figure 2.4: *Constructing a chordal $G'$ from the chordal graph $P_4$ by adding a $K_5$, connecting its vertices pairwise to $G$. Adding the (blue) auxiliary vertices are necessary to preserve chordality.*

*Proof.* We will give a reduction from DOMINATING SET on chordal graphs. Given $G = (V, E)$ with vertex set $V = \{v_1, ... v_n\}$, we construct a chordal graph $G'$ as described below:

1. Add one complete graph $K_{n+1}$ consisting of the vertices $\{k_1, ..., k_n, u\}$ and an edge $\{v_i, k_i\}$ to each vertex $v_i \in V$ of $G$. One vertex of the complete subgraph is not connected to any $v \in V$. Denote it as $u$.

2. Add one additional vertex $t$ and connect it with $u$ vie the edge $\{u, t\}$.

3. For all vertices $v_i \in V$ in $G$, add a new edge $\{n, k_i\}$ for all neighbors $n \in N(v_i)$.

An example reduction on the graph $P_4$ is shown in section 2.3.4.

**Corollary 2.3.2.** *$N(v_i) \in G$ forms a clique iff $N(v_i)$ forms a clique in $G'$*

*Proof.* Asuming that $N(v_i)$ forms a clique in G, we show that it also forms a clique in $G'$ by induction over the number of neighbors $z = abs(N(v_i))$ in G.

- $z = 0$: Holds trivially as we do not have a neighbor in G and in G' the connected $k_i$ forms a $P_1$, hence a clique.

- $z = z + 1$:

  By IH, we already know that all neighbors $n_1, ..., n_z$ form a clique together with their vertices in $k_i$. As $k_{z+1}, v_{z+1} \in N(v_i)$ now also in G', we show that $N(v_i)$ still forms clique in G'.

  Let $k_i$ be the vertex that was connected with $n_i$ during step 1. All we have to show is that $v_{z+1}$ and $k_{z+1}$ extend our previous clique, hence are fully connected with $N(v_i)$.
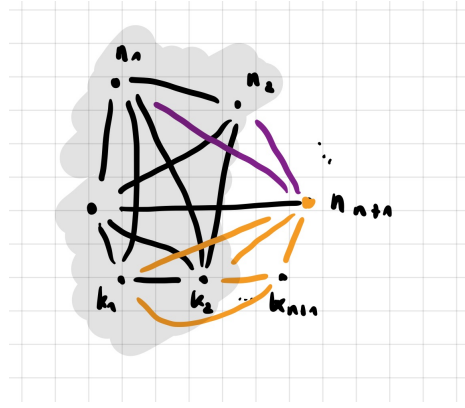
Figure 2.5: Induction Step

$v_{z+1}$ connects to $N(v_i)$ in G by assumption. By our construction, there exists an edge to $k_1, ..., k_z$, because we add an edge $(n_{z+1}, k_i)$ if there is an edge from $(n_{z+1}, n_i)$. (See fig 2.5)

$k_{z+1}$ form a complete subgraph with the other $k_i$ and is connected to all $n_i$ by construction because the edge $(n_{z+1}, n_i)$ exists.

Therefore, $N(v_i)$ will also form a clique in $G'$.

On the other side, if $N(v_i)$ forms a clique in G', the vertices of $N(v_i)$ in G form an induced subgraph of G', hence preserving the clique.

∎

**Corollary 2.3.3.** *G is Chordal iff G' is chordal.*

*Proof.* ⇒: Assume G chordal. Then exists a total elimination order $o = (v_1, ..., v_n)$ in G where removing $v_j$ sequentially returns cliques in $N(v_i)$. Define $o' = (v_1, ..., v_n, k_1, ..., k_n, u, t)$. Applying corollary 2.3.2 states that $(v_1, ...v_n)$ always gives cliques in G and according to corollary 2.3.2 also in G'. As the rest is directly part of a clique in G' by definition with an additional vertex of degree 1, o' is a total elimination order for $G'$, hence G' chordal. ⇐: Holds as o' is always a total elimination order in G' and removing the complete subgraph $K_{n+1}$ and $u$ gives a total elimination order in G. ∎

**Corollary 2.3.4.** *G has a Dominating Set of size k iff G' has a sds of size $k + 1$*

*Proof.* Assume a ds $D$ of size $k$ in G. $D \cup \{u\}$ is an sds in $G'$ of size $k + 1$, because $u$ dominates $t$ and for each $v \in DS : d(v, u) \leq 2$.

Contrary, assume an sds $SD$ in $G'$. To dominate $t$, $u \in SD$ must hold, hence already dominating the complete subgraph $K_{n+1}$. If a vertex $k_i \in SD$, we exchange it with $v_i$ not losing the domination property. Taking $D = SD - \{u\}$ gives our desired ds of size $k$. ∎

As this reduction runs in FPT time and the parameter is only bounded by a function of k, this is an FPT reduction. As Dominating Set on Chordal Graphs is $w[2] - hard$, so is Semitotal Dominating Set on Chordal Graphs.

$\square$

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