Tight Lower Bounds on Graph Embedding Problems*

Marek Cygan¹, Fedor V. Fomin^{2,5}, Alexander Golovnev^{3,5}, Alexander S. Kulikov⁵, Ivan Mihajlin⁴, Jakub Pachocki⁶, and Arkadiusz Socała¹

¹Institute of Informatics, University of Warsaw, Poland
²University of Bergen, Norway
³New York University, USA
⁴University of California—San Diego, USA
⁵St. Petersburg Department of Steklov Institute of Mathematics of the Russian Academy of Sciences, Russia
⁶Carnegie Mellon University, USA

Abstract

We prove that unless the Exponential Time Hypothesis (ETH) fails, deciding if there is a homomorphism from graph G to graph H cannot be done in time $|V(H)|^{o(|V(G)|)}$. We also show an exponential-time reduction from Graph Homomorphism to Subgraph Isomorphism. This rules out (subject to ETH) a possibility of $|V(H)|^{o(|V(H)|)}$ -time algorithm deciding if graph G is a subgraph of H. For both problems our lower bounds asymptotically match the running time of brute-force algorithms trying all possible mappings of one graph into another. Thus, our work closes the gap in the known complexity of these fundamental problems.

Moreover, as a consequence of our reductions conditional lower bounds follow for other related problems such as Locally Injective Homomorphism, Graph Minors, Topological Graph Minors, Minimum Distortion Embedding and Quadratic Assignment Problem.

1 Introduction

We establish tight conditional lower bounds on the complexity of several fundamental graph embedding problems including Graph Homomorphism, Subgraph Isomorphism, Graph Minor, Topological Graph Minor, and Minimum Distortion Embedding. For given undirected graphs G and H, all these problems can be solved in time $n^{\mathcal{O}(n)}$ by a brute-force algorithm that tries all possible embeddings of G into H, where n is the number of vertices in G and H. We show that unless the Exponential Time Hypothesis (ETH) fails, the running time $n^{\mathcal{O}(n)}$ is unavoidable. This resolves a number of open problems about graph embeddings that can be found in the literature. We start by defining embedding problems and providing for each of the problems a brief overview of the related previous results.

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Graph Homomorphism A homomorphism $G \to H$ from an undirected graph G to an undirected graph H is a mapping from the vertex set G to that of H such that the image of every edge of G is an edge of H. In other words, there is $G \to H$ if and only if there exists a mapping $g: V(G) \to V(H)$, such that for every edge $uv \in E(G)$, we have $g(u)g(v) \in E(H)$. Then the Graph Homomorphism problem HOM(G, H) is defined as follows.

Graph Homomorphism

Input: Undirected graphs G and H.

Task: Decide whether there is a homomorphism $G \to H$.

Many combinatorial structures in G, for example cliques, independent sets, and proper vertex colorings, may be viewed as graph homomorphisms to a particular graph H, see the book of Hell and Nešetřil [2004] for a thorough introduction to the topic. It is well-known that COLORING is a special case of graph homomorphism. More precisely, a graph G can be colored with at most h colors if and only if $G \to K_h$, where K_h is a complete graph on h vertices. Due to this, very often in the literature HOM(G, H), when h = |V(H)|, is referred as H-coloring of G. It was shown by Feder and Vardi [1998] that Constraint Satisfaction Problem (CSP) can be interpreted as a homomorphism problem on relational structures, and thus Graph Homomorphism encompasses a large family of problems generalizing Coloring but less general than CSP.

Hell and Nešetřil showed that for every fixed simple graph H, the problem whether there exists a homomorphism from G to H is solvable in polynomial time if H is bipartite, and NP-complete if H is not bipartite [Hell and Nešetřil 1990]. Since then, algorithms for and the complexity of graph homomorphisms (and homomorphisms between other discrete structures) have been studied intensively [Austrin 2010; Barto et al. 2008; Grohe 2007; Marx 2010; Raghavendra 2008].

There are two different ways graph homomorphisms are used to extract useful information about graphs. Let us consider two homomorphisms, from a "small" graph F into a "large" graph G and from a "large" graph G into a "small" graph H, which can be represented by the following formula (here we borrow the intuitive description from the book of Lovász [2012])

$$F \to G \to H$$
.

Then "left-homomorphisms" from various small graphs F into G are useful to study the local structure of G. For example, if F is a triangle, then the number of "left-homomorphisms" from F into G is the number of triangles in graph G. This type of information is closely related to sampling, and we refer to the book of Lovász [2012] which provides many applications of homomorphisms. "Right-homomorphisms" into "small" different graphs H are related to global properties of graph G.

The trivial brute-force algorithm solving "left-homomorphism" from an f-vertex graph F into an n-vertex graph G runs in time $2^{\mathcal{O}(f\log n)}$: we try all possible vertex subsets of G of size at most f, which is $n^{\mathcal{O}(f)}$ and then for each subset try all possible f^f mappings into it from F. Interestingly, this naïve algorithm is asymptotically optimal. Indeed, as it was shown by Chen et al. [2006], assuming Exponential Time Hypothesis, there is no $g(k)n^{o(k)}$ time algorithm deciding if an input n-vertex graph G contains a clique of size at least k, for any computable function g. Since this is a very special case of Graph Homomorphism with F being a clique of size k, the result of Chen et al. rules out algorithms for Graph Homomorphism of running time $g(f)2^{o(f\log n)}$, from F to G, when the number of vertices f in F is significantly smaller than the number of vertices n in G.

The interest in "right-homomorphisms" is due to the recent developments in the area of exact exponential algorithms for Coloring and 2-CSP (CSP where all constraints have arity at most 2)

problems. The area of exact exponential algorithms is about solving intractable problems significantly faster than the trivial exhaustive search, though still in exponential time [Fomin and Kratsch 2010]. For example, as for Graph Homomorphism, a naïve brute-force algorithm for coloring an n-vertex graph G in h colors is to try for every vertex a possible color, resulting in the running time $\mathcal{O}^*(h^n) = 2^{\mathcal{O}(n \log h)}$. Since h can be of order $\Omega(n)$, the brute-force algorithm computing the chromatic number runs in time $2^{\mathcal{O}(n \log n)}$. It was already observed in 1970s by Lawler [1976] that the brute-force for the Coloring problem can be beaten by making use of dynamic programming over maximal independent sets resulting in single-exponential running time $\mathcal{O}^*((1+\sqrt[3]{3})^n) = \mathcal{O}(2.45^n)$. Almost 30 years later Björklund et al. [2009] succeeded to reduce the running time to $\mathcal{O}^*(2^n)$. And as we observed already, for H-coloring, the brute-force algorithm solving H-coloring runs in time $2^{\mathcal{O}(n \log h)}$. In spite of all the similarities between graph coloring and homomorphism, no substantially faster algorithm was known and it was an open question in the area of exact algorithms if there is a single-exponential algorithm solving H-coloring in time $2^{\mathcal{O}(n+h)}$ [Fomin et al. 2007; Rzażewski 2014; Wahlström 2010; 2011], see also [Fomin and Kratsch 2010, Chapter 12].

On the other hand, Graph Homomorphism is a special case of 2-CSP with n variables and domain of size h. It was shown by Traxler [2008] that unless the Exponential Time Hypothesis fails, there is no algorithm solving 2-CSP with n variables and domain of size h in time $h^{o(n)} = 2^{o(n \log h)}$. This excludes (up to ETH) the existence of a single-exponential c^n time algorithm for some constant c > 1 for 2-CSP.

Another interesting variant of Graph Homomorphism is related to graph labelings. A homomorphism $f: G \to H$ is called *locally injective* if for every vertex $u \in V(G)$, its neighborhood is mapped injectively into the neighborhood of f(u) in H, i.e., if every two vertices with a common neighbor in G are mapped onto distinct vertices in H.

LOCALLY INJECTIVE GRAPH HOMOMORPHISM

Input: Undirected graphs G and H.

Task: Decide whether there is a locally injective homomorphism $G \to H$.

As graph homomorphism generalizes graph coloring, locally injective graph homomorphism can be seen as a generalization of graph distance constrained labelings. An L(2,1)-labeling of a graph G is a mapping from V(G) into the nonnegative integers such that the labels assigned to vertices at distance 2 are different while labels assigned to adjacent vertices differ by at least 2. This problem was studied intensively in combinatorics and algorithms, see, e.g., Griggs and Yeh [1992] and Fiala et al. [2008]. Fiala and Kratochvíl suggested the following generalization of L(2,1)-labeling, we refer [Fiala and Kratochvíl 2008] for the survey. For graphs G and H, an H(2,1)-labeling is a mapping $f:V(G)\to V(H)$ such that for every pair of distinct adjacent vertices $u,v\in V(G)$, images f(u) f(v) are distinct and nonadjacent in H. Moreover, if the distance between u and v in G is two, then $f(u)\neq f(v)$. It is easy to see that a graph G has an L(2,1)-labeling with maximum label at most k if and only if there is an H(2,1)-labeling for H being a k-vertex path. Then the following is known, see for example [Fiala and Kratochvíl 2008]: there is an H(2,1)-labeling of a graph G if and only if there is a locally injective homomorphism from G to the complement of H.

Several single-exponential algorithms for L(2,1)-labeling can be found in the literature, the

 $^{{}^1\}mathcal{O}^*(\cdot)$ hides polynomial factors in the input length. Most of the algorithms considered in this paper take graphs G and H as an input. By saying that such an algorithm has a running time $\mathcal{O}^*(f(G,H))$, we mean that the running time is upper bounded by $(|V(G)| + |E(G)| + |V(H)| + |E(H)|)^{\mathcal{O}(1)} \cdot f(G,H)$.

most recent algorithm is due to Junosza-Szaniawski et al. [2013] which runs in time $\mathcal{O}(2.6488^n)$. For H(2,1)-labeling, or equivalently for locally injective homomorphisms, single-exponential algorithms were known only for special cases when the maximum degree of H is bounded [Havet et al. 2011] or when the bandwidth of the complement of H is bounded [Rzażewski 2014].

Subgraph Isomorphism We say that an undirected G is a *subgraph* of H if one can remove some edges and vertices of H, so that what remains is isomorphic to G. In other words, G is a subgraph of H if and only if there exists an injective mapping $g: V(G) \to V(H)$, such that for each edge $uv \in E(G)$, $g(u)g(v) \in E(H)$. We define

Subgraph Isomorphism

Input: Undirected graphs G and H.

Task: Decide whether G is a subgraph of H.

SUBGRAPH ISOMORPHISM is an important and very general problem. Several flagship graph problems can be viewed as instances of SUBGRAPH ISOMORPHISM:

- Hamiltonicity(G): Is C_n (a cycle with n vertices) a subgraph of G?
- CLIQUE(G, k): Is K_k a subgraph of G?
- 3-COLORING(G): Is G a subgraph of $K_{n,n,n}$, a tripartite graph with n vertices in each of its three independent sets?
- BANDWIDTH(G, k): Is G a subgraph of P_n^k (a k-th power of an n-vertex path)?

All of the mentioned problems are NP-complete, and the best known algorithms for all the listed special cases work in exponential time. In fact, all those problems are well-studied from the exact exponential algorithms perspective [Beigel and Eppstein 2005; Björklund 2014; Bourgeois et al. 2012; Cygan and Pilipczuk 2012; Feige 2000; Held and Karp 1962; Lawler 1976; Robson 1986; Tarjan and Trojanowski 1977], where the goal is to obtain an algorithm of running time $\mathcal{O}(c^n)$ for the smallest possible value of c. Furthermore, the Subgraph Isomorphism problem was very extensively studied from the viewpoint of fixed parameter tractability, see [Marx and Pilipczuk 2014] for a discussion of 19 different possible parametrizations. All the mentioned special cases of Subgraph Isomorphism admit $\mathcal{O}(c^n)$ time algorithms, by using either branching, inclusion-exclusion principle, or dynamic programming. On the other hand, a simple exhaustive search for the Subgraph—runs in $2^{\mathcal{O}(n \log n)}$ time, where n is the total number of vertices of the host graph and pattern graph.

Therefore, a natural question is whether Subgraph Isomorphism admits an $\mathcal{O}(c^n)$ time algorithm. This was repeatedly posed as an open problem [Cygan et al. 2014; Amini et al. 2012; Fomin et al. 2008; Husfeldt et al. 2013]. In particular, in the monograph of Fomin and Kratsch [2010] the existence of $\mathcal{O}(c^n)$ time algorithm for Subgraph Isomorphism was put among the few questions in the open problems section.

SUBGRAPH ISOMORPHISM is a special case of QUADRATIC ASSIGNMENT PROBLEM, which is

QUADRATIC ASSIGNMENT PROBLEM (QAP)

Input: $n \times n$ matrices $A = (a_{ij})$ and $B = (b_{ij})$ with real entries.

Task: Find a permutation π minimizing $\sum_{i=1}^{n} \sum_{j=1}^{n} a_{\pi(i)\pi(j)} b_{ij}$.

Indeed, G is a subgraph of H if and only if for the instance of QAP with A and B being adjacency matrices of G and the complement of H the optimum value is 0^2 . Problem 7.6 in the influential survey of Woeginger on exact algorithms [Woeginger 2003] is to prove that QAP cannot be solved in time $\mathcal{O}(c^n)$ for any fixed value c (under some reasonable assumption).

Graph Minor For a graph G and an edge $uv \in G$, we define the operation of contracting edge uv as follows: we delete vertices u and v from G, and add a new vertex w_{uv} adjacent to all vertices that u or v was adjacent to in G. We say that a graph G is a minor of H, if G can be obtained from some subgraph of H by a series of edge contractions. Equivalently, we may say that G is a minor of H if G can be obtained from H itself by a series of edge deletions, edge contractions and vertex deletions.

GRAPH MINOR

Input: Undirected graphs G and H. **Task:** Decide whether G is a minor of H.

Graph Minor is a fundamental problem in graph theory and graph algorithms. By the theorem of Robertson and Seymour [1995], there exists a computable function f and an algorithm that, for given graphs G and H, checks in time $f(G)|V(H)|^3$ whether G is a minor of H. However, when the size of the graph G is not a constant, nothing beyond a brute-force algorithm trying all possible partitions of a vertex set of H was known.

Related notion of graph embedding is the notion of topological minor. We say that a graph G is a subdivision of a graph H if G can be obtained from H by contracting only edges incident with vertices of degree two. In other words, G is obtained from H by replacing edges with paths. A graph G is called a *topological minor* of a graph H if a subdivision of G is isomorphic to a subgraph of H.

TOPOLOGICAL GRAPH MINOR

Input: Undirected graphs G and H.

Task: Decide whether G is a topological minor of H.

Lingas and Wahlen [2009] gave an algorithm of running time $\mathcal{O}^*(\binom{n}{p}p!2^{n-p})$ solving TOPOLOGICAL GRAPH MINOR for *n*-vertex graph H and *p*-vertex graph G.

Minimum Distortion Embedding Given an undirected connected graph G with the vertex set V(G) and the edge set E(G), the graph metric of G is $M(G) = (V(G), D_G)$, where the distance function D_G is the shortest path distance between u and v for every pair of vertices $u, v \in V(G)$. Given a graph metric M and another metric space M' with distance functions D and D', a mapping $f: M \to M'$ is called an embedding of M into M'. The mapping f is non-contracting, if for every pair of points p, q in M, $D(p, q) \leq D'(f(p), f(q))$. The distortion of embedding f is the minimum number d_f such that $D(p, q) \cdot d_f \geq D'(f(p), f(q))$. We define

MINIMUM DISTORTION EMBEDDING

Input: Undirected graphs G and H.

Task: Find a non-contracting embedding of G into H of minimum distortion.

²If G has smaller number of vertices than H, then it should be first padded with isolated vertices to make the number of vertices in both graphs equal.

Most of exact algorithms for MINIMUM DISTORTION EMBEDDING deal with a special case when the host graph H is a path or a tree of bounded degree [Bădoiu et al. 2005a;b; Cygan and Pilipczuk 2012; Fellows et al. 2013; Fomin et al. 2011; Kenyon et al. 2009]. In particular, an optimal-distortion embedding into a line can be found in time $2^{\mathcal{O}(n)}$ [Cygan and Pilipczuk 2012; Fomin et al. 2011].

Our results. In this paper we show that from the algorithmic perspective, the behavior of "right-homomorphism" is, unfortunately, much closer to 2-CSP than to COLORING. This result will also imply similar lower bounds for many other graph embedding and containment problems. All lower bounds obtained in this paper are conditional, they hold unless the Exponential Time Hypothesis [Impagliazzo and Paturi 2001; Impagliazzo et al. 2001b] fails. ETH is an established assumption; many interesting lower bounds have been found under this hypothesis (see [Cygan et al. 2015; Lokshtanov et al. 2011] for surveys). We formulate ETH in the next section.

The first main result of this paper is the following theorem, which excludes (up to ETH) resolvability of HOM(G, H) in time $2^{o(n \log h)}$, thus resolving the open qestion from [Fomin et al. 2007; Rzażewski 2014; Wahlström 2010; 2011].

Theorem 1. Unless ETH fails, for any constant d > 0 there exists a constant c = c(d) > 0 such that for any non-decreasing function $3 \le h(n) \le n^d$, there is no algorithm solving Graph Homomorphism from an n-vertex graph G to a graph H with at most h(n) vertices in time

$$\mathcal{O}(2^{cn\log h(n)}). \tag{1}$$

Let us remark that in order to obtain more general results, in all lower bounds proven in this paper we assume implicitly that the number h of vertices of the graph H is a function of the number n of the vertices of the graph G. At the same time, to exclude some pathological cases we assume that the function h(n) is "reasonable" meaning that it is non-decreasing and time-constructible.

With a tiny modification, the proof of Theorem 1 can be adapted to show a similar lower bound for LOCALLY INJECTIVE GRAPH HOMOMORPHISM.

Theorem 2. Unless ETH fails, for any constant d > 0 there exists a constant c = c(d) > 0 such that for any non-decreasing function $3 \le h(n) \le n^d$, there is no algorithm deciding if there is a locally injective homomorphism from an n-vertex graph G to a graph H with at most h(n) vertices in time $\mathcal{O}(2^{cn \log h(n)})$.

The second main result of this paper is about Subgraph Isomorphism, resolving the open question asked in [Cygan et al. 2014; Amini et al. 2012; Fomin et al. 2008; Husfeldt et al. 2013; Fomin and Kratsch 2010].

Theorem 3. Unless ETH fails, there is no algorithm solving Subgraph Isomorphism for graphs G and H in time $2^{o(n \log n)}$, where n = |V(G)| + |V(H)|.

Theorem 3 implies that QAP cannot be solved in time $2^{o(n \log n)}$ unless ETH fails and hence provides the answer to the open problem of Woeginger [2003].

An important feature of our proof is that it rules out solvability of Subgraph Isomorphism in time $2^{o(n \log n)}$ even for the special case when |V(G)| = |V(H)| = n. Since in this special case a graph G is a (topological) minor of H if and only if G is a subgraph of H. Thus the case of Graph Minor and Topological Graph Minor when |V(G)| = |V(H)| = n cannot be resolved in time $2^{o(n \log n)}$ as well. Similar arguments work for various modifications of Graph Minor like Shallow Graph Minor, etc.

To see how the bound on Subgraph Isomorphism yields the bound on Minimum Distortion Embedding, we observe that an n-vertex graph G admits a non-contracting embedding of distortion 1 into an n-vertex graph H if and only if H is a subgraph of G.

Methods To establish lower bounds for graph homomorhisms, we proceed in two steps. First we obtain lower bounds for List Graph Homomorphism by reducing it to the 3-coloring problem on graphs of bounded degree. More precisely, for a given graph G with vertices of small degrees, we construct an instance (G', H') of List Graph Homomorphism, such that G is 3-colorable if and only if there exists a list homomorphism from G' to H'. Moreover, our construction guarantees that a "fast" algorithm for list homomorphism implies an algorithm for 3-coloring violating ETH. The reduction is based on a "grouping" technique, however, to do the required grouping we need a trick exploiting the condition that G has a bounded maximum vertex degree and thus can be colored in a bounded number of colors in polynomial time. In the second step of reductions we proceed from list homomorphisms to normal homomorphisms. Here we need specific gadgets with a property that any homomorphism from such a graph to itself preserves an order of its specific structures.

The remaining part of the paper is organized as follows. Section 2 contains all necessary definitions. In Section 3 we give technical lemmata and reductions which are used to prove lower bounds for the Graph Homomorphism problem in Section 4.1 and for the Subgraph Isomorphism in Section 4.2. We conclude with some open problems in Section 5.

2 Preliminaries

Graphs We consider simple undirected graphs, where V(G) denotes the set of vertices and E(G) denotes the set of edges of a graph G. For a given subset S of V(G), G[S] denotes the subgraph of G induced by S, and G - S denotes the graph $G[V(G) \setminus S]$. A vertex set S of G is an independent set if G[S] is a graph with no edges, and S is a clique if G[S] is a complete graph. The set of neighbors of a vertex v in G is denoted by $N_G(v)$, and the set of neighbors of a vertex set S is $N_G(S) = \bigcup_{v \in S} N_G(v) \setminus S$. By $N_G[S]$ we denote the closed neighborhood of the set S, i.e., the set S together with all its neighbors: $N_G[S] = S \cup N_G(S)$. For an integer S, we use S to denote the set of integers S integers S in S in S in S in S integers S in S in S in S in S in S integers S in S

The complete graph on k vertices is denoted by K_k . A coloring of a graph G is a function assigning a color to each vertex of G such that adjacent vertices have different colors. A k-coloring of a graph uses at most k colors, and the chromatic number $\chi(G)$ is the smallest number of colors in a coloring of G. By Brook's theorem, for any connected graph G with maximum degree G is at most G unless G is a complete graph, in which case the chromatic number is G is at most G unless G is a graph can be found in polynomial time by a straightforward greedy algorithm.

Throughout the paper we implicitly assume that there is a total order on the set of vertices of a given graph. This allows us to treat a k-coloring of a n-vertex graph simply as a vector in $[k]^n$.

Let G be an n-vertex graph, $1 \leq r \leq n$ be an integer, and $V(G) = B_1 \sqcup B_2 \sqcup \ldots \sqcup B_{\lceil \frac{n}{r} \rceil}$ be a partition of the set of vertices of G. Then the grouping of G with respect to the partition $V(G) = B_1 \sqcup B_2 \sqcup \ldots \sqcup B_{\lceil \frac{n}{r} \rceil}$ is a graph G_r with vertices $B_1, \ldots, B_{\lceil \frac{n}{r} \rceil}$ such that B_i and B_j are adjacent if and only if there exist $u \in B_i$ and $v \in B_j$ such that $uv \in E(G)$. To distinguish vertices of the graphs G and G_r , the vertices of G_r will be called buckets.

For a graph G, its square G^2 has the same set of vertices as G and $uv \in E(G^2)$ if and only if there is a path of length at most 2 between u and v in G (thus, $E(G) \subseteq E(G^2)$). It is easy to see that if the degree of G is less than Δ then the degree of G^2 is less than Δ^2 and hence a Δ^2 -coloring of G^2 can be easily found.

Homomorphisms and list homomorphisms Let G and H be graphs. A mapping $\varphi: V(G) \to V(H)$ is a homomorphism if for every edge $uv \in E(G)$ its image $\varphi(u)\varphi(v) \in E(H)$. If there exists a homomorphism from G to H, we often write $G \to H$. The Graph Homomorphism problem HOM(G, H) asks whether or not $G \to H$.

Assume that for each vertex v of G we are given a list $\mathcal{L}(v) \subseteq V(H)$. A list homomorphism of G to H, also known as a list H-coloring of G, with respect to the lists \mathcal{L} , is a homomorphism $\varphi: V(G) \to V(H)$, such that $\varphi(v) \in \mathcal{L}(v)$ for all $v \in V(G)$. The LIST GRAPH HOMOMORPHISM problem LIST-HOM(G, H) asks whether or not graph G with lists \mathcal{L} admits a list homomorphism to H with respect to \mathcal{L} .

Exponential Time Hypothesis Our lower bounds are based on a well-known complexity hypothesis formulated by Impagliazzo et al. [2001a].

Exponential Time Hypothesis (ETH): There is a constant q > 0 such that 3-CNF-SAT with n variables and m clauses cannot be solved in time $2^{qn}(n+m)^{\mathcal{O}(1)}$.

This hypothesis is widely applied in the theory of exact exponential algorithms, we refer to [Cygan et al. 2015; Lokshtanov et al. 2013] for an overview of ETH and its implications.

In this paper we use the following well-known application of ETH with respect to 3-Coloring (see, e.g., Theorem 3.2 in [Lokshtanov et al. 2013], and Exercise 7.27 in [Sipser 2005]). The 3-Coloring problem is the problem to decide whether the given graph can be properly colored in 3 colors.

Proposition 1. Unless ETH fails, there exists a constant q > 0 such that 3-Coloring on n-vertex graphs of average degree four cannot be solved in time $\mathcal{O}^*(2^{qn})$.

It is well known that 3-Coloring remains NP-complete on graphs of maximum vertex degree four. Moreover, the classical reduction, see e.g. [Garey and Johnson 1979], allows for a given n-vertex graph G to construct a graph G' with maximum vertex degree at most four and $|V(G')| = \mathcal{O}(|E(G)|)$ such that G is 3-colorable if and only if G' is. Thus Proposition 1 implies the following (folklore) lemma which will be used in our proofs.

Lemma 1. Unless ETH fails, there exists a constant q > 0 such that there is no algorithm solving 3-Coloring on n-vertex graphs of maximum degree four in time $\mathcal{O}^*(2^{qn})$.

3 Auxiliary Lemmata

In this section we provide reductions and auxiliary lemmata about colorings which will be used to prove lower bounds for Graph Homomorphism and Subgraph Isomorphism.

3.1 Balanced Colorings

In the following we show how to construct a specific "balanced" coloring of a graph in polynomial time. Let G be a graph of constant maximum degree. The coloring of G we want to construct should satisfy three properties. First, it should be a proper coloring of G^2 . Then the size of each color class should be bounded as well as the number of edges between vertices from different color classes. More precisely, we prove the following lemma.

Lemma 2. For any constant d, there exist constants $\alpha, \beta, \tau > 1$ and a polynomial time algorithm that for a given graph G on n vertices of maximum degree d and an integer $\tau \leq L \leq \frac{n(d^2-1)}{2d^2(d^2+1)}$, finds a coloring $c: V(G) \to [L]$ satisfying the following properties:

- 1. The coloring c is a proper coloring of G^2 .
- 2. There are only a few vertices of each color: for all $i \in [L]$,

$$|c^{-1}(i)| \le \left\lceil \alpha \cdot \frac{n}{L} \right\rceil \,. \tag{2}$$

3. There are only a few edges of G between each pair of colors: For all $i \neq j \in [L]$, we have

$$k_{i,j} := |\{uv \in E(G) : c(u) = i, c(v) = j\}| \le K_{i,j} := \left\lceil \beta \cdot \frac{\min\{|c^{-1}(i)|, |c^{-1}(j)|\}}{L} \right\rceil.$$

Proof. The algorithm starts by constructing greedily an independent set I of G^2 of size $\left\lceil \frac{n}{d^2+1} \right\rceil$. Since the maximum vertex degree of G^2 does not exceed d^2 , this is always possible. We construct a partial coloring of G^2 by coloring the vertices of I in L colors such that the obtained coloring is a balanced coloring of $G^2[I]$, meaning that the number of vertices of each color is $\lfloor |I|/L \rfloor$ or $\lceil |I|/L \rceil$. Since I is an independent set in G^2 , such a coloring can be easily constructed in polynomial time. In the obtained partial equitable coloring, we have that for every $i \in [L]$

$$|c^{-1}(i)| \ge \left| \frac{n}{L(d^2+1)} \right| \ge \frac{n}{2Ld^2}$$
 (3)

(recall that $L \leq \frac{n(d^2-1)}{2d^2(d^2+1)}$). Let us note that the obtained precoloring of G^2 clearly satisfies the first and the third conditions of the lemma. Since the size of every $c^{-1}(i)$, $i \in [L]$, does not exceed $|c^{-1}(i)| \leq \left\lceil \frac{n}{L} \right\rceil$, the second condition of the lemma also holds for every $\alpha \geq 1$.

We extend the precoloring of G^2 to the required coloring by the following greedy procedure: We select an arbitrary uncolored vertex v and color it by a color from [L] such that the new partial coloring also satisfies the three conditions of the lemma. In what follows, we prove that such a greedy choice of a color is always possible.

Coloring of a vertex v with a color i can be forbidden only because it breaks one of the three conditions. Let us count, how many colors can be forbidden for v by each of the three constraints.

- 1. Vertex v has at most d^2 neighbors in G^2 , so the first constraint forbids at most d^2 colors.
- 2. The second constraint forbids all the colors that are "fully packed" already. The number of such colors is at most $\frac{n}{\left(\frac{\alpha n}{L}\right)} = \frac{L}{\alpha}$.

3. To estimate the number of colors forbidden by the third condition, we go through all the neighbors of v. A neighbor $u \in N_G(v)$ forbids a color i if coloring v by i exceeds the allowed bound on $k_{i,c(u)}$. Hence to estimate the number of such forbidden colors i (for every fixed vertex u) we need to estimate how many values of $k_{i,c(u)}$ can reach the allowed upper bound $K_{i,c(u)}$. We have that

$$\begin{split} \left| \left\{ i \colon k_{i,c(u)} = K_{i,c(u)} \right\} \right| & \overset{\text{by (3)}}{\leq} \left| \left\{ i \colon k_{i,c(u)} \geq \frac{\beta n}{2L^2 d^2} \right\} \right| = \left| \left\{ i \colon k_{i,c(u)} \cdot \frac{2L^2 d^2}{\beta n} \geq 1 \right\} \right| \\ & \leq \sum_{i \in [L]} k_{i,c(u)} \cdot \frac{2L^2 d^2}{\beta n}. \end{split}$$

The number of edges between vertices of the same color c(u) and all other vertices of the graph does not exceed the cardinality of the color class c(u) times d. Thus we have

$$\begin{split} \sum_{i \in [L]} k_{i,c(u)} \cdot \frac{2L^2 d^2}{\beta n} &\leq d|c^{-1}(c(u))| \cdot \frac{2L^2 d^2}{\beta n} \overset{\text{by } (2)}{\leq} d \left\lceil \frac{\alpha n}{L} \right\rceil \cdot \frac{2L^2 d^2}{\beta n} \\ &\leq d \frac{2\alpha n}{L} \cdot \frac{2L^2 d^2}{\beta n} = \frac{4\alpha L d^3}{\beta} \,. \end{split}$$

where the last inequality is due to $\alpha > 1$ and $L \leq n$.

Therefore,

$$|\{i \colon k_{i,c(u)} = K_{i,c(u)}\}| \le \frac{4\alpha Ld^3}{\beta}.$$

Since the degree of v in G does not exceed d, we have that the number of colors forbidden by the third constraint is at most $\frac{4\alpha L d^4}{\beta}$.

Thus, the total number of colors forbidden by all the three constraints for the vertex v is at most

$$d^2 + \frac{L}{\alpha} + \frac{4\alpha L d^4}{\beta}$$
.

By taking sufficiently large constants α , β , and τ , say $\alpha = 4$, $\beta = 16\alpha^2d^4$, and $\tau = \frac{16(d^2+1)}{11}$, we guarantee that this expression does not exceed L-1 for every $L \geq \tau$. Therefore, there always exists a vacant color for the vertex v which concludes the proof.

Now with help of Lemma 2, we describe a way to construct a specific grouping of a graph. The properties of such groupings are crucial for the final reduction.

Lemma 3. For any constant d, there exists a constant $\lambda = \lambda(d)$ and a polynomial time algorithm that for a given graph G on n vertices of maximum degree d and an integer $r \leq \sqrt{\frac{n}{2\lambda}}$, finds a grouping \tilde{G} of G and a coloring \tilde{c} : $V(\tilde{G}) \to [\lambda r]$ such that

1. The number of buckets of \tilde{G} is

$$|V(\tilde{G})| \le \frac{|V(G)|}{r};$$

- 2. The coloring \tilde{c} is a proper coloring of \tilde{G}^2 ;
- 3. Each bucket $B \in V(\tilde{G})$ is an independent set in G, i.e. for every $u, v \in B$, $uv \notin E(G)$;
- 4. For every pair of buckets $B_1, B_2 \in V(\tilde{G})$ there is at most one edge between them in G, i.e.

$$|\{uv \in E(G): u \in B_1, v \in B_2\}| \le 1.$$

Proof. Let $\beta = \beta(d)$ be a constant provided by Lemma 2 and let $L = \lambda r$ for $\lambda = \lambda(d) = 2d\beta$. Let also c be a coloring of G in L colors provided by Lemma 2. We want to construct a grouping \tilde{G} of G such that for all buckets $B \in V(\tilde{G})$ and all $u \neq v \in B$,

$$c(u) = c(v) \text{ and } c(u') \neq c(v')$$

for all $u' \in N_G(u), v' \in N_G(v)$. (4)

In other words, all vertices of the same bucket are of the same color while any two neighbors of such two vertices are of different colors.

For each color $i \in [L]$, we introduce an auxiliary constraint graph F_i . The vertex set of F_i is $V(F_i) = c^{-1}(i)$ and its edge set is

$$E(F_i) = \{uv : \exists u' \in N_G(u), v' \in N_G(v), c(u') = c(v')\}.$$

In our construction, each bucket of \tilde{G} will be an independent set in some F_i . Note that this will immediately imply (4). The degree of any vertex $v \in V(F_i)$ is at most

$$\deg_{F_i}(v) \le \sum_{v' \in N_G(v)} (K_{c(v), c(v')} - 1) \le d \left(\left\lceil \frac{\beta |c^{-1}(v)|}{L} \right\rceil - 1 \right) \le \frac{d\beta |V(F_i)|}{L} = \frac{|V(F_i)|}{2r}.$$

This means that the greedy algorithm finds a proper coloring of each F_i in at most $\frac{|V(F_i)|}{2r} + 1$ colors, which splits each F_i in at most $\frac{|V(F_i)|}{2r} + 1$ independent sets. We create a separate bucket of \tilde{G} from each independent set of each F_i . Now we show that the four conditions from the lemma statement hold.

1. For the first property, the number of independent sets in each F_i is at most $\frac{|V(F_i)|}{2r} + 1$. Thus the number of buckets in \tilde{G} is

$$|V(\tilde{G})| \le \sum_{i \in [L]} \left(\frac{|V(F_i)|}{2r} + 1 \right) = \sum_{i \in [L]} \left(\frac{|c^{-1}(i)|}{2r} + 1 \right) = \frac{n}{2r} + L \le \frac{n}{r}.$$

since $L = \lambda r$ and $2\lambda r^2 \le n$.

- 2. For the second property, by Lemma 2, the coloring c is proper in G^2 . We can convert c to a coloring $\tilde{c}: V(\tilde{G}) \to [\lambda r]$ by assigning each bucket the color of its vertices (all of them have the same color). The resulting coloring \tilde{c} is a proper coloring of \tilde{G}^2 by (4) and the fact that c is a proper in G^2 .
- 3. All buckets of \tilde{G} are monochromatic with respect to c, thus, each bucket $B \in V(\tilde{G})$ is an independent set in G and the third property holds.
- 4. Finally, by (4), there is at most one edge in G between vertices corresponding to any pair of buckets in \tilde{G} .

Thus, the constructed grouping and its coloring satisfy all conditions of the lemma. \Box

3.2 Reductions

This section constitutes the main technical part of the paper and contains all the necessary reductions used in the lower bounds proofs. Using these reductions as building blocks the lower bounds follow from careful calculations. The general pipeline is as follows. To prove a lower bound, we take a graph G of maximum degree four that needs to be 3-colored and construct an equisatisfiable instance (G', H') of List Graph Homomorphism using Lemma 4. We then use Lemma 5 to transform (G', H') into an equisatisfiable instance (G'', H'') of Graph Homomorphism. Thus, an algorithm checking whether there exists a homomorphism from G'' to H'' can be used to check whether the initial graph G can be 3-colored. At the same time we know a lower bound for 3-Coloring under ETH (Lemma 1). This gives us a lower bound for Graph Homomorphism, we show an exponential-time reduction from Graph Homomorphism to Subgraph Isomorphism.

Lemma 4 (3-Coloring \to List Graph Homomorphism). There exists an algorithm that takes as input a graph G on n vertices of maximum degree d that needs to be 3-colored and an integer $r = o(\sqrt{n})$ and finds an equisatisfiable instance (G', H') of LIST-HOM, where $|V(G')| \le n/r$ and $|V(H')| \le \gamma(d)^r$, where $\gamma(d)$ is a function of the graph degree. The running time of the algorithm is polynomial in n and the size of the output graphs.

Proof. Constructing the graph G'. Let G' be the grouping of G and $c: V(G') \to [L]$ be the coloring provided by Lemma 3 where $L = \lambda(d)r$. To distinguish colorings of G and G', we call c(B), for a bucket $B \in V(G')$, a label of G. Consider a bucket $G \in V(G')$, i.e., a subset of vertices of G, and a label $G \in [L]$. From item 2 of Lemma 3 we know that $G \in [L]$ is a proper coloring of G'. This, in particular, means that there is at most one $G' \in N_{G'}(G)$ such that $G' \in [L]$ is a Moreover, if such $G' \in [L]$ is then, by item 4 of Lemma 3, there exists a unique $G \in [L]$ and unique $G' \in [L]$ is such that $G' \in [L]$ is allows us to define the following mapping $G \in [L]$ is $G \in [L]$ if such $G' \in [L]$ if $G \in [L]$ if $G \in [L]$ if $G \in [L]$ is an oneighbor $G' \in [L]$ if $G \in [L]$ if $G \in [L]$ is a neighbor outside of its bucket (it cannot have a neighbor in its own bucket as buckets are independent), $G \in [L]$.

Constructing the graph H'. We now define a redundant encoding of a 3-coloring of a bucket $B \in V(G')$. Namely, let $\mu_B : (f : B \to \{1,2,3\}) \to \{0,1,2,3\}^L$. That is, for a 3-coloring $f : B \to \{1,2,3\}$ of B, μ_B is a vector v of length L. For $i \in [L]$, by v[i] we denote the i-th component of v. The value of v[i] is defined as follows: if $\phi_B(i) = 0$ then v[i] = 0, otherwise $v[i] = f(\phi_B(i))$. In other words, for a given bucket B and a 3-coloring f of its vertices, for each possible label $i \in [L]$, μ_B is the color of the vertex $u \in B$ that has a neighbor in a bucket with label i, and 0 if there is no such vertex u.

We are now ready to construct the graph H'. The set of vertices of H' is defined as follows:

$$V(H') = \left\{(R,l) \colon R \in \left\{0,1,2,3\right\}^L \text{ and } l \in [L]\right\},$$

i.e., a vertex of H' is an encoding of a 3-coloring of a bucket and a label of a bucket. The list constraints of this instance of LIST GRAPH HOMOMORPHISM are defined as follows: a bucket $B \in V(G')$ is allowed to be mapped to $(R,l) \in V(H')$ if and only if l = c(B) and there is a 3-coloring f of B such that $\mu_B(f) = R$. Informally, two vertices in V(H') are joined by an edge if they define two consistent 3-colorings. Formally, $(R_1, l_1)(R_2, l_2) \in E(H')$ if and only if $R_1[l_2] \neq R_2[l_1]$. Note that $|V(G')| \leq n/r$ by Lemma 3 and $|V(H')| \leq 4^L \cdot L \leq 4^L \cdot 2^L = 8^{\lambda(d)r} = \gamma(d)^r$ for $\gamma(d) = 8^{\lambda(d)}$.

Running time of the reduction. The reduction clearly takes time polynomial in the size of input and output.

Correctness of the reduction. It remains to show that G is 3-colorable if and only if (G', H') is a yes-instance of LIST GRAPH HOMOMORPHISM.

Assume that G is 3-colorable and take a proper 3-coloring g of G. It defines a homomorphism from G' to H' in a natural way: $B \in V(G')$ is mapped to $(\mu_B(g|_B), c(B))$, where $g|_B$ is the function g with its domain restricted to B. Each list constraint is satisfied by definition. To show that each edge is mapped to an edge, consider an edge $BB' \in E(G')$. Then, by item 4 of Lemma 3 there is a unique edge $uu' \in E(G)$ such that $u \in B, u' \in B'$. Note that B and B' are mapped to vertices (R, l) and (R', l') such that R[l'] = g(u) and R'[l] = g(u'). Since g is a proper 3-coloring of G, $g(u) \neq g(u')$. This, in turn, means that $(R, l)(R', l') \in E(H')$ and hence the edge BB' is mapped to this edge in H'.

For the reverse direction, consider a homomorphism $h : G' \to H'$. For each bucket $B \in V(G')$, h(B) defines a proper 3-coloring of B. Together, they define a 3-coloring g of G and we need to show that g is proper. Assume, to the contrary, that there is an edge $uu' \in E(G)$ such that g(u) = g(u'). By item 3 of Lemma 3, u and u' belong to different buckets $B, B' \in V(G')$. By the definition of grouping, $BB' \in E(G')$. Since h is a homomorphism, $(R, l)(R', l') := h(B)h(B') \in E(H')$. At the same time, R[l'] = g(u) = g(u') = R'[l] which contradicts the fact that (R, l)(R', l') is an edge in H'.

Lemma 5 (LIST GRAPH HOMOMORPHISM \rightarrow GRAPH HOMOMORPHISM). There is a polynomial-time algorithm that from an instance (G,H) of LIST-HOM where |V(G)|=n, |V(H)|=h constructs an equisatisfiable instance (G',H') of HOM where $|V(G')| \leq n + \Delta$, $|V(H')| \leq \Delta$ for $\Delta = 25h^2$.

Proof. Preparations. We start with a simple 6-vertex gadget D' consisting of a 5-cycle together with an apex vertex adjacent to all the vertices of the cycle, see Fig. 1.

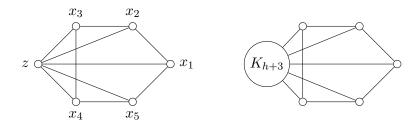


Figure 1: The graphs D' and D. The encircled clique K_{h+3} is the canonical clique of D. An edge from a clique to a vertex of a cycle means that each vertex of the clique is joined to this vertex.

An important property of D' is that for each homomorphism $\phi \colon D' \to D'$ and $i \in [5]$,

$$\phi(z) = z$$
 and $\phi(z) \neq \phi(x_i)$.

In words, z is always mapped to z and nothing else is mapped to z. Indeed, because the vertex z is adjacent to all the remaining vertices of D', we have that $\phi(z) \neq \phi(x_i)$. By the same reason, we have that for every $i \in [5]$, $\phi(x_i) \in N_{D'}(\phi(z))$. But for every x_i its open neighborhood $N_{D'}(x_i)$

induces a bipartite graph. On the other hand, the chromatic number of the cycle $C = x_1x_2x_3x_4x_5$ is three, and thus it cannot be mapped by ϕ to $N_{D'}(x_i)$ for any $i \in [5]$. Therefore, $\phi(z) = z$.

In order for the $\phi(z) = z$ argument to work in a bigger graph, we replace z by a clique K_{h+3} of size h+3, called the *canonical clique* of the gadget. The obtained graph with (h+3)+5 vertices is denoted as D (see Fig. 1).

Let D_0, \ldots, D_k be k+1 copies of the graph D. We join those k+1 graphs isomorphic to D to construct a larger gadget T_k as follows (see Fig. 2). For each $i \in [k]$, we select an arbitrary vertex from the canonical clique of D_i , denote this vertex as z_i , and identify it with one arbitrary vertex of D_{i-1} which does not belong to the canonical clique of D_{i-1} , i.e., with a vertex of the 5-cycle and connect every vertex of K_{h+3} to all neighbors of z in the subsequent block. We also mark one of K_{h+3} 's vertices as z and connect it to the left of it, see Fig. 2. Denote the new graph by $T_{k,h+3}$. Oberve that each D_i is a block of T_k and we call D_i the ith block of T_k . Note that two consecutive blocks D_{i-1} and D_i have exactly one common vertex, namely z_i .

The reason we are using those canonical cliques instead of single vertices in the construction of T_k is that those canonical cliques are big enough to behave as anchors. That is, we will prove that canonical cliques can only be mapped to themselves and not to other parts of the graph, in particular, for each $i \in [k]$ and homomorphism $\phi: T_k \to T_k$, $\phi(z_i) = z_i$.

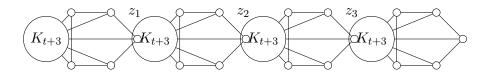


Figure 2: The gadget T_k for k = 3.

Constructing G'. Let A_h be a graph consisting of a matching with h edges $\{a_1b_1, \ldots, a_hb_h\}$. Then the graph G' consists of a copy of G, a copy of T_h , and a copy of A_h with the following additional edges: the vertex z_i from the ith block of T_h is adjacent to the vertices a_i and b_i . Also we add edges from G to A_h : for a vertex $g_i \in G$ we add an edge g_ia_j for every j, and an edge g_ib_j if $j \notin \mathcal{L}(i)$ (see Fig. 3). The number of vertices in G' is at most $n + 2h + (h + 1)(h + 3 + 5) \leq n + (h + 1)(h + 11) \leq n + 25h^2$.

Constructing H'. The graph H' is constructed similarly. It consists of a copy of H, a copy of T_h , and a copy of A_h . For every i we add edges $z_i a_i$ and $z_i b_i$ as before. Also, each vertex i of H is adjacent to all the vertices from A_h except for b_i (see Fig. 4). The number of vertices in H' is at most $h + 2h + (h + 1)(h + 3 + 5) \le (h + 1)(h + 11) \le 25h^2$.

Correctness. We now turn to prove that the instance (G, H) of LIST-HOM is equisatisfiable to an instance (G', H') of HOM.

Claim 1. Any homomorphism ϕ from G' to H' maps T_h into T_h .

Proof of the claim. No pair of vertices of the same clique of T_h is mapped to the same vertex in H', because H' has no self-loops. Therefore, canonical cliques from T_h are mapped to some cliques from T_h , as H' has no more cliques of size h+3. The remaining vertices of T_h have at least h+3 neighbors from canonical cliques, therefore they must be mapped to vertices from T_h .

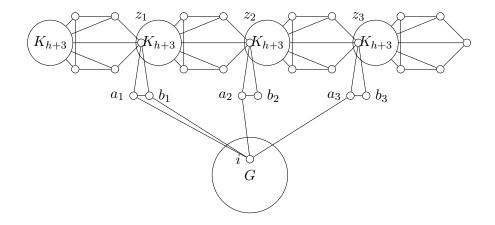


Figure 3: The graph G'. A vertex $i \in V(G)$ is connected to b_j if and only if $j \notin \mathcal{L}(i)$, where $\mathcal{L}(i)$ is the list associated with the vertex $i \in V(G)$.

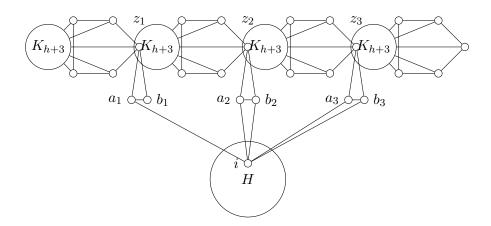


Figure 4: The graph H'. A vertex $i \in V(H)$ is connected to all a_i 's and all b_i 's except for b_i .

Claim 2. Any homomorphism ϕ from G' to H' bijectively maps T_h to T_h so that the order of z's is preserved, i.e., for each $i \in [k]$, $\phi(z_i) = z_i$.

Proof of the claim. 1. Every canonical clique is mapped to a canonical clique. First note that a canonical clique is mapped into one block. Indeed, there are no vertices outside of a block that are connected to more than one vertex of the block. Assume, to the contrary, that same canonical clique is mapped to one block but not to a canonical clique. Then its image has to contain one or two vertices of the 5-cycle from that block. If the image contains only one vertex of the 5-cycle, then the image of the 5-cycle has at most 3 vertices: one vertex from the canonical clique K_{h+3} , two neighbors of the vertex from the 5-cycle (because all the vertices of the image of the 5-cycle must be connected to all the vertices of the image of the clique). Note that these three vertices do not form a triangle, therefore the 5-cycle cannot be mapped to them. If the image of the clique contains two vertices outside of the canonical clique, then for

the same reason the image of the 5-cycle must contain only two vertices, which is not possible. This analysis shows that every canonical clique K_{h+3} must be mapped to a canonical clique K_{h+3} .

- 2. Every block is mapped to a block. We already know that every canonical clique is mapped to a canonical clique. The 5-cycle from the same block must be mapped to the corresponding 5-cycle, because it is the only image that contains a closed walk of odd length and every vertex of which is connected to the clique (recall that the images of the canonical clique and the 5-cycle do not intersect, since their preimages are joined by edges). Note that since the canonical clique and the cycle are mapped to themselves, z_i has to be mapped to some z_j .
- 3. If D_i is mapped to D_j , then D_{i+1} is mapped to D_{j+1} . The cycle from D_i shares a vertex with the canonical clique from D_{i+1} , therefore if D_i is mapped to D_j , then D_{i+1} can only be mapped to D_{j+1} or D_j . However, D_{i+1} cannot be mapped into the same block as D_i . Indeed, in this case the canonical clique of D_{i+1} would be mapped to the canonical clique of D_j , but we already know that z_{i+1} is mapped to the 5-cycle of D_j . Therefore, D_i and D_{i+1} must be mapped in consecutive blocks.

The above proves that for every $i \in \{0, ..., k\}$, D_i is mapped to D_i , which implies that any homomorphism preserves the order of z's.

Claim 3. Any homomorphism ϕ from G' to H' maps A_h to A_h so that a_ib_i is mapped to a_ib_i .

Proof of the claim. Every pair a_ib_i is connected to $z_i \in T_h$, so it can be mapped either to a_ib_i or to some vertices of T_h . But in the latter case it would not have paths of length 2 to all other pairs a_ib_i .

Claim 4. Any homomorphism ϕ from G' to H' maps G to H.

Proof of the claim. Assume, to the contrary, that a vertex $g \in V(G)$ is mapped to a vertex $v \in V(T_h)$ or a vertex $a \in V(A_h)$. The vertex g is adjacent to at least h vertices from A_h , but v and a are adjacent to at most 2 vertices from A_h (recall that by the previous claim every a_ib_i is mapped to a_ib_i).

Now we show that the two instances are equisatisfiable. Let ϕ be a list homomorphism from G to H. We show that its natural extension ϕ' mapping T_h to T_h and A_h to A_h is a correct homomorphism from G' to H'. This is non-trivial only for edges of G' from G to A_h . Consider an edge from a vertex i of G to a vertex b_j . The presence of this edge means that i is not mapped to j by ϕ . Recall that the b_j is mapped by ϕ to b_j . This means that the considered edge in G' is mapped to an edge in H' by ϕ' .

For the reverse direction, let ϕ' be a homomorphism from G' to H'. We show that its natural projection is a list homomorphism from G to H. Since ϕ' maps G to H (by Claim 4) it is enough to check that all list constrains are satisfied. For this, consider a vertex i from G and assume that $j \notin \mathcal{L}(i)$. Then ϕ' does not map i to j as otherwise there would be no image for one of the edges $g_i a_j$ or $g_i b_j$, where g_i is the ith vertex of G.

Running time of the reduction. The reduction clearly takes time polynomial in the input length.

4 Lower Bounds

4.1 Graph Homomorphism

We are ready to prove our main result about graph homomorphisms, i.e., Theorem 1.

Theorem 4 (Theorem 1 restated). Let G be an n-vertex graph and H be a graph with at most h := h(n) vertices. Unless ETH fails, for any constant $D \ge 1$ there exists a constant c = c(D) > 0 such that for any non-decreasing function $0 \le n$, there is no $O(h^{cn})$ time algorithm deciding whether there is a homomorphism from G to H.

Proof. The outline of the proof of the theorem is as follows. Assuming that there is a "fast" algorithm for Graph Homomorphism, we show that there is also a "fast" algorithm solving List Graph Homomorphism, which, in turn, implies "fast" algorithm for 3-Coloring on degree 4 graphs, contradicting ETH. In what follows, we specify what we mean by "fast".

Let $h_0 = 25^2$. If $h(n) < h_0$ for all values of n, then an algorithm with running time $\mathcal{O}(h^{cn})$ would solve 3-Coloring in time $\mathcal{O}(h_0^{cn}) = \mathcal{O}(2^{cn \log h_0})$ (recall that $h(n) \geq 3$). Therefore, by choosing a small enough constant c such that $c \log h_0 < q$, we arrive to a contradiction with Lemma 1.

From now on we assume that $h(n) \geq h_0$ for large enough values of n. Let $c = \frac{q}{8D\log\gamma}$, where q is the constant from Lemma 1, and $\gamma := \gamma(4)$ is the constant from Lemma 4. For the sake of contradiction, let us assume that there exists an algorithm \mathcal{A} deciding whether $G \to H$ in time $\mathcal{O}(h^{cn}) = \mathcal{O}(2^{cn\log h})$, where |V(G)| = n, |V(H)| = h := h(n). Now we show how to solve 3-coloring on n'-vertex graphs of maximum degree four in time $2^{qn'}$, which would contradict Lemma 1.

Let G' be an n'-vertex graph of maximum degree four that needs to be 3-colored. Let $r = \frac{\log h}{4D \log \gamma}$ and $n = \frac{2n'}{r}$. Using Lemma 4 (note that $r = o(\sqrt{n'})$ as required) we construct an instance (G_1, H_1) of List Graph Homomorphism that is satisfiable if and only if the initial graph G' is 3-colorable, and $|V(G_1)| \leq \frac{n'}{r}, |V(H_1)| \leq \gamma^r$. By Lemma 5, this instance is equisatisfiable to an instance (G, H) of Graph Homomorphism where $|V(H)| < 25\gamma^{2r} = 25h^{\frac{1}{2D}} \leq h$ (since $D \geq 1$ and $h(n) \geq h_0$), and

$$|V(G)| \le \frac{n'}{r} + 25\gamma^{2r} \le \frac{n}{2} + 25h^{\frac{1}{2D}} \le \frac{n}{2} + 25\sqrt{n} \le n$$

(for sufficiently large values of n).

Now, in order to solve 3-coloring for G', we construct an instance (G,H) with $|V(G)| \leq n$ and $|V(H)| \leq h$ of Graph Homomorphism and invoke the algorithm $\mathcal A$ on this instance. The running time of $\mathcal A$ is

$$\mathcal{O}(2^{cn\log h}) = \mathcal{O}(2^{\frac{2cn'}{r}\log h}) = \mathcal{O}(2^{\frac{2cn'}{\log h}\log h}) = \mathcal{O}(2^{8cDn'\log \gamma}) = \mathcal{O}(2^{8cDn'\log \gamma}) = \mathcal{O}(2^{qn'})$$

and hence we can find a 3-coloring of G' in time $\mathcal{O}(2^{qn'})$, which contradicts ETH (see Lemma 1). \square

Theorem 5 (Theorem 2 restated). Let G be an n-vertex graph G and H be a graph with at most h := h(n) vertices. Unless ETH fails, for any constant $D \ge 1$ there exists a constant c = c(D) > 0 such that for any non-decreasing function $0 \le n$, there is no $0 \le n$ time algorithm deciding whether there is a locally injective homomorphism from $0 \le n$.

Proof. The proof is almost identical to the proof of Theorem 4.

Let us observe that in the reduction in Lemma 4, in graph G', we take a coloring (in the proof we refer to such coloring as to labeling) of the square of G'. Thus for every bucket v of G', all its

neighbors are labeled by different colors. The way we construct the lists, only buckets with the same labels can be mapped to the same vertex of H'. Thus for every vertex v of G', no pair of its neighbors can be mapped to the same vertex. Hence every list homomorphism from G' to H' is locally injective. Therefore the result of Lemma 4 holds for locally injective list homomorphisms as well and we obtain the following lemma.

Lemma 6. There exists an algorithm that takes as input a graph G on n vertices of maximum degree d that needs to be 3-colored and an integer $r = o(\sqrt{n})$ and finds an equisatisfiable instance (G', H') of LOCALLY INJECTIVE GRAPH HOMOMORPHISM, where $|V(G')| \leq n/r$ and $|V(H')| \leq \gamma(d)^r$, where $\gamma(d)$ is a function of the graph degree. The running time of the algorithm is polynomial in n and the size of the output graphs.

In the reduction of Lemma 5, we established that every homomorphism from G' to H' maps T_h to T_h and A_h to A_h so that a_ib_i is mapped to a_ib_i . Thus for vertices of these structures, every homomorphism is locally injective. By Claim 4, any homomorphism ϕ from G' to H' maps G to H. Therefore there is a locally injective homomorphism from G' to H' if and only if there is a locally injective list homomorphism from G to G to

4.2 Subgraph Isomorphism

To prove a lower bound for Subgraph Isomorphism we need a reduction, which given an instance of Graph Homomorphism produces a single exponential number of instances of Subgraph Isomorphism. Even though from the perspective of polynomial time algorithms such a reduction gives no implication in terms of which problem is harder, in our setting it is enough to obtain a lower bound for Subgraph Isomorphism.

Theorem 6. Given an instance (G, H) of Graph Homomorphism one can in $poly(n)2^n$ time create 2^n instances of Subgraph Isomorphism with n vertices, where n = |V(G)| + |V(H)|, such that (G, H) is a yes-instance if and only if at least one of the created instances of Subgraph Isomorphism is a yes-instance.

Proof. Let (G, H) be an instance of Graph Homomorphism and let n = V(G) + V(H). Note that any homomorphism h from G to H can be associated with some sequence of non-negative numbers $(|h^{-1}(v)|)_{v \in V(H)}$, being the numbers of vertices of G mapped to particular vertices of H. The sum of the numbers in such a sequence equals exactly |V(G)|. As the number of such sequences is $\binom{V(G)+V(H)-1}{V(H)-1} \le 2^n$, we can enumerate all such sequences in time 2^n poly(n). For each such sequence $(a_v)_{v \in V(H)}$ we create a new instance (G', H') of Subgraph Isomorphism, where the pattern graph remains the same, i.e., G' = G, and in the host graph H' each vertex of $v \in V(H)$ is replicated exactly a_v times (possibly zero). Observe that |V(H')| = |V(G')|.

We claim that G admits a homomorphism to H if and only if for some sequence $(a_v)_{v \in V(H)}$ the graph G' is a subgraph of H'. First, assume that G admits a homomorphism h to H. Consider the instance (G', H') created for the sequence $a_v = |h^{-1}(v)|$ and observe that we can create a bijection $h': V(G') \to V(H')$ by assigning $v \in V(G')$ to its private copy of h(v). As h is a homomorphism, so is h', and as h' is at the same time a bijection, we infer that G' is a subgraph of H'.

On the other hand if for some sequence $(a_v)_{v \in V(H)}$ the constructed graph G' is a subgraph of H', then projecting the witnessing injection $g: V(G') \to V(H')$ so that g'(v) is defined as the

prototype of the copy g(v) gives a homomorphism from G to H, as copies of each $v \in V(H)$ form independent sets in H'.

Combining Theorem 4 with Theorem 6, we immediately obtain the following lower bound.

Theorem 7. Unless ETH fails, there exists a constant c > 0 such that there is no algorithm deciding whether a given n-vertex graph G contains a subgraph isomorphic to a given n-vertex graph H in time $\mathcal{O}(n^{cn})$.

5 Conclusion and Open Problems

In this work we resolved a number of questions about exact exponential algorithms. Our lower bounds suggest several directions for further research.

"Fine-grained" dichotomy The classical results of Hell and Nešetřil [1990] establishes the following dichotomy for Graph Homomorphism subject to $P \neq NP$: For every fixed simple graph H, the problem whether there exists a homomorphism from G to H is solvable in polynomial time if and only if H is bipartite. Is there anything similar to that in the world of exponential algorithms for HOM(G, H)?

More precisely, for graph classes \mathcal{G} and \mathcal{H} we denote by $HOM(\mathcal{G}, \mathcal{H})$ the restriction of the graph homomorphism problem to input graphs $G \in \mathcal{G}$ and $H \in \mathcal{H}$. If \mathcal{G} or \mathcal{H} is the class of all graphs then we use the placeholder ' \bot ' instead of a letter. Thus the result of Hell-Nešetřil states that unless $P \neq NP$, $HOM(_, \mathcal{H})$ is in P if and only if \mathcal{H} is a class of bipartite graphs.

Now we know that solving $\mathrm{HOM}(\underline{\ },\underline{\ })$ with input graphs G and H in time $|V(H)|^{o(|V(G)|)}$ would refute ETH. On the other hand, when $\mathcal H$ is the class of graphs consisting of complete graphs, $\mathrm{HOM}(\underline{\ },\mathcal H)$ is equivalent to computing the chromatic number of G and thus is solvable in time $\mathcal O(2^{|V(G)|})$ [Björklund et al. 2009]. More generally, let $\mathcal H$ be a graph class such that for some constant t, either the clique-width or the maximum vertex degree of the core of every graph in $\mathcal H$ is at most t. Wahlström [2010] have shown that in this case $\mathrm{HOM}(\underline{\ },\mathcal H)$ is solvable in single-exponential time $\mathcal O(f(t)^{|V(G)|}) = 2^{\mathcal O(|V(G)|)}$, where f is some function of $\mathcal H$ only. Is it possible to characterize (up to some complexity assumption) graph classes $\mathcal H$, where $\mathrm{HOM}(\underline{\ },\mathcal H)$ is solvable in single-exponential time?

What about the fine-grained complexity of Graph Homomorphism for $HOM(\mathcal{G}, \bot)$ and $HOM(\mathcal{G}, \mathcal{H})$? Of course, similar questions are interesting for Subgraph Isomorphism, as well as for counting versions of Graph Homomorphism and Subgraph Isomorphism.

Some concrete problems Are the following problems solvable in single-exponential time?

- SUBGRAPH ISOMORPHISM with instance (G, H) when the maximum vertex degree of G is 3. (When degree of G does not exceed 2, the problem is solvable in single-exponential time, see e.g. [Held and Karp 1962].)
- \bullet Deciding if graph G can be obtained from graph H only by edge-contractions.
- Deciding if graph G is an immersion of graph H.
- Deciding if G is a minor of a graph H for the special situation when G is a clique.

• Finding a minimum distortion embedding into a cycle. We remark that embedding in a path can be done in time $2^{\mathcal{O}(|V(G)|)}$ [Cygan and Pilipczuk 2012; Fomin et al. 2011].

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