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Approximating Highly Inapproximable Problems on Graphs of Bounded Twin-Width

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Abstract

For any \( \varepsilon > 0 \), we give a polynomial-time \( n^{\varepsilon} \)-approximation algorithm for MAX INDEPENDENT SET in graphs of bounded twin-width given with an \( O(1) \)-sequence. This result is derived from the following time-approximation trade-off: We establish an \( O(1)^{2q-1} \)-approximation algorithm running in time \( \exp(O(n^{1-q})) \), for every integer \( q \geq 0 \). Guided by the same framework, we obtain similar approximation algorithms for MIN COLORING and MAX INDUCED MATCHING. In general graphs, all these problems are known to be highly inapproximable: for any \( \varepsilon > 0 \), a polynomial-time \( n^{1-\varepsilon} \)-approximation for any of them would imply that \( P=NP \) [Håstad, FOCS ‘96; Zuckerman, ToC ‘07; Chalermsook et al., SODA ‘13]. We generalize the algorithms for MAX INDEPENDENT SET and MAX INDUCED MATCHING to the independent (induced) packing of any fixed connected graph \( H \).

In contrast, we show that such approximation guarantees on graphs of bounded twin-width given with an \( O(1) \)-sequence are very unlikely for MIN INDEPENDENT DOMINATING SET, and somewhat unlikely for LONGEST PATH and LONGEST INDUCED PATH. Regarding the existence of better approximation algorithms, there is a (very) light evidence that the obtained approximation factor of \( n^{1-\varepsilon} \) for MAX INDEPENDENT SET may be best possible. This is the first in-depth study of the approximability of problems in graphs of bounded twin-width. Prior to this paper, essentially the only such result was a polynomial-time \( O(1) \)-approximation algorithm for MIN DOMINATING SET [Bonnet et al., ICALP ‘21].

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1 Introduction

Twin-width is a graph parameter introduced by Bonnet, Kim, Thomassé, and Watrigant [10]. Its definition involves the notions of trigraphs and of contraction sequences. A trigraph is a graph with two types of edges: black (regular) edges and red (error) edges. A (vertex) contraction consists of merging two (non-necessarily adjacent) vertices, say, \( u, v \) into a vertex \( w \), and keeping every edge \( wz \) black if and only if \( uz \) and \( vz \) were previously black edges. The other edges incident to \( w \) become red (if not already), and the rest of the trigraph remains the same. A contraction sequence of an \( n \)-vertex trigraph \( G \) is a sequence of trigraphs

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3 In this introduction, we might implicitly use \( n \) to denote the number of vertices, and \( m \), the number of edges of the graph at hand.
A \( d \)-sequence is a contraction sequence in which every vertex of every trigraph has at most \( d \) red edges incident to it. The twin-width of \( G \), denoted by \( \text{tww}(G) \), is then the minimum integer \( d \) such that \( G \) admits a \( d \)-sequence. Figure 1 gives an example of a graph with a 2-sequence, i.e., of twin-width at most 2. Twin-width can be naturally extended to matrices (with unordered \([10]\) or ordered \([8]\) row and column sets) over a finite alphabet, and thus to binary structures.

An equivalent viewpoint that will be somewhat more convenient is to consider a \( d \)-sequence as a sequence of partitions \( P_n := \{\{v\} : v \in V(G)\}, P_{n-1}, \ldots, P_1 := \{V(G)\} \) of \( V(G) \), such that for every integer \( 1 \leq i \leq n-1 \), \( P_i \) has \( i \) parts and is obtained by merging two parts of \( P_{i+1} \) into one. Now the red degree of a part \( P \in P_i \) is the number of other parts \( Q \in P_i \) such that there is in \( G \) at least one edge and at least one non-edge between \( P \) and \( Q \).

A \( d \)-sequence is such that no part of no partition of the sequence has red degree more than \( d \). In that case the maximum red degree of each partition is at most \( d \). And we similarly get the twin-width of \( G \) as the minimum integer \( d \) such that \( G \) admits a (partition) \( d \)-sequence. The quotient trigraph \( G/P_1 \) is the trigraph \( G_{i_1} \), if the (contraction) \( d \)-sequence \( G_n, \ldots, G_1 \) and the (partition) \( d \)-sequence \( P_n, \ldots, P_1 \) correspond.

Classes of binary structures with bounded twin-width include graph classes with bounded treewidth, and more generally bounded clique-width, proper minor-closed classes, posets with antichains of bounded size, strict subclasses of permutation graphs, as well as \( \Omega(\log n) \)-subdivisions of \( n \)-vertex graphs \([10]\), and some classes of (bounded-degree) expanders \([5]\).

A notable variety of geometrically defined graph classes have bounded twin-width such as map graphs, bounded-degree string graphs \([10]\), classes with bounded queue number or bounded stack number \([5]\), segment graphs with no \( K_{t,t} \) subgraph, visibility graphs of 1.5D terrains without large half-graphs, visibility graphs of simple polygons without large independent sets \([4]\).

For every class \( \mathcal{C} \) mentioned so far, \( O(1) \)-sequences can be computed in polynomial time\(^2\) on members of \( \mathcal{C} \). For classes of binary structures including a binary relation interpreted as a linear order on the domain (called ordered binary structures), there is a fixed-parameter approximation algorithm for twin-width \([8]\). More precisely, given a graph \( G \) and an integer \( k \), there are computable functions \( f \) and \( g \) such that one can output an \( f(k) \)-sequence of \( G \) or correctly report that \( \text{tww}(G) > k \) in time \( g(k)n^{O(1)} \). Such an approximation algorithm is currently missing for classes of general (not necessarily ordered) binary structures, and in particular for the class of all graphs. We also observe that deciding if the twin-width of a graph is at most 4 is an NP-complete task \([3]\).

We will therefore assume that the input graph is given with a \( d \)-sequence, and treat \( d \) as a constant (or that the input comes from any of the above-mentioned classes). Thus far, this is the adopted setting when designing faster algorithms on bounded twin-width.

\(^2\) Admittedly, for the geometric classes, a representation is (at least partially) needed.
graphs [10, 7, 33, 30, 19]. From the inception of twin-width [10] –actually already from the seminal work of Guillemot and Marx [21]– it was clear that structures wherein this invariant is bounded may often allow the design of parameterized algorithms. More concretely, it was shown [10] that, on graphs $G$ given with a $d$-sequence, model checking a first-order sentence $\varphi$ is fixed-parameter tractable—it can be solved in time $f(d, \varphi) \cdot n^c$, the special cases of, say, $k$-INDEPENDENT SET or $k$-DOMINATING SET admit single-exponential parameterized algorithms [7], an effective data structure almost linear in $n$ can support constant-time edge queries [33], the triangles of $G$ can be counted in time $O(d^2 n + m)$ [30].

So far, however, the connection between having bounded twin-width and enjoying enhanced approximation factors was tenuous. The only such result concerned MIN DOMINATING SET, known to be inapproximable in polynomial-time within factor $(1 − o(1)) \ln n$ unless P=NP [16], but yet admits a constant-approximation on graphs of bounded twin-width given with an $O(1)$-sequence [7]. We start filling this gap by designing approximation algorithms on graphs of bounded twin-width given with an $O(1)$-sequence for notably MAX INDEPENDENT SET (MIS, for short), MAX INDUCED MATCHING, and COLORING. Getting better approximation algorithms for MIS and COLORING in that particular scenario was raised as an open problem [7]. Before we describe our results and elaborate on the developed techniques, let us briefly present the notorious inapproximability of these problems in general graphs.

MIS and COLORING are NP-hard [20], and very inapproximable: for every $\varepsilon > 0$, it is NP-hard to approximate these problems within ratio $n^{1-\varepsilon}$ [23, 34]. The same was shown to hold for MAX INDUCED MATCHING [13]. Besides, there is only little room to improve over the brute-force algorithm in $2^{O(n)}$: Unless the Exponential Time Hypothesis$^3$ [25] (ETH) fails, no algorithm can solve MIS in time $2^{o(n)}$ [26] (nor the other two problems). For any $r$ (possibly a function of $n$) WMIS can be $r$-approximated in time $2^{O(n/r)}$ [15, 12]. Bansal et al. [2] essentially shaved a $\log^2 r$ factor to the latter exponent. It is known though that polynomial shavings are unlikely. Chalermsook et al. [14] showed that, for any $\varepsilon > 0$ and sufficiently large $r$ (again $r$ can be function of $n$), an $r$-approximation for MIS and MAX INDUCED MATCHING cannot take time $2^{O(n^{1-\varepsilon} / r^{1+\varepsilon})}$, unless the ETH fails. For instance, investing time $2^{O(\sqrt{n})}$, one cannot hope for significantly better than a $\sqrt{n}$-approximation.

Contributions and techniques

Our starting point is a constant-approximation algorithm for MIS running in time $2^{O(\sqrt{n})}$ when presented with an $O(1)$-sequence, which is very unlikely to hold in general graphs by the result of Chalermsook et al. [14].

**Theorem 1.** On $n$-vertex graphs given with a $d$-sequence MAX INDEPENDENT SET can be $O_d(1)$-approximated in time $2^{O_d(\sqrt{n})}$.

Our algorithm builds upon the functional equivalence between twin-width and the so-called versatile twin-width [5]. We defer the reader to Section 2 for a formal definition of versatile twin-width. For our purpose, one only needs to know the following useful consequence of that equivalence. From a $d'$-sequence of $G$, we can compute in polynomial time another partition sequence $P_1, \ldots, P_n$ of $G$ of width $d := f(d')$, for some computable function $f$, such that for every integer $1 \leq i \leq n$, all the $i$ parts of $P_i$ have size at most $d \cdot \frac{n}{i}$. Even if some parts of $P_i$ can be very small, this partition is balanced in the sense that no part can be larger than $d$ times the part size in a perfectly balanced partition. Of importance to us is $P_i[\sqrt{n}]$ when the

$^3$ That is, the assumption that there is a $\delta > 0$ such that $n$-variable 3-SAT cannot be solved in time $\delta^n$. 
number of parts (\(\lfloor \sqrt{n} \rfloor \)) and the size of a larger part in the partition (at most \(d \frac{n}{\lfloor \sqrt{n} \rfloor} \approx d\sqrt{n}\)) are somewhat level.

We can then properly color the red graph (made by the red edges on the vertex set \(\mathcal{P}_{\lfloor \sqrt{n} \rfloor}\)) with \(d + 1\) colors. Any color class \(X\) is a subset of parts of \(\mathcal{P}_{\lfloor \sqrt{n} \rfloor}\) such that between two parts there are either all edges (black edge) or no edge at all (non-edge). In graph-theoretic terms, the subgraph \(G_X\) of \(G\) induced by all the vertices of \(X\) have a simple modular decomposition: a partition of at most \(\sqrt{n}\) modules each of size at most \(d\sqrt{n}\). It is thus routine to compute a largest independent set of \(G_X\) essentially in time exponential in the maximum between the number of modules and the maximum size of a module, that is, in at most \(d\sqrt{n}\). As one color class \(X^*\) contains more than a \(\frac{1}{d\sqrt{n}}\) fraction of the optimum, we get our \(d + 1\)-approximation when computing a largest independent set of \(G_X\) . Figure 2 on page 13 serves as a visual summary of what we described so far.

The next step is to substitute recursive calls of our approximation algorithm to exact exponential algorithms on induced subgraphs of size \(O_d(\sqrt{n})\). Following this inductive process at depth \(q = 2, 3, 4, \ldots\), we degrade the approximation ratio to \((d + 1)^3, (d + 1)^7, (d + 1)^{15}\), etc. but meanwhile we boost the running time to \(2^{O_d(n^{1/4})}, 2^{O_d(n^{1/7})}, 2^{O_d(n^{1/15})}\), etc. In effect we show by induction that:

\[ \textbf{Theorem 2.} \] On \(n\)-vertex graphs given with a \(d\)-sequence \textsc{Max Independent Set} has an \(O_d(1)^{2^{d-1}}\)-approximation algorithm running in time \(2^{O_d(n^{2^{d-1}})}\), for every integer \(q \geq 0\).

The following polynomial-time algorithm is a corollary of Theorem 2 choosing \(q = O_d, (\log \log n)\).

\[ \textbf{Theorem 3.} \] For every \(\varepsilon > 0\), \textsc{Max Independent Set} can be \(n^\varepsilon\)-approximated in polynomial-time \(O_{d,\varepsilon}(1) \cdot \log^{O_d(1)} n \cdot n^O(1)\) on \(n\)-vertex graphs given with a \(d\)-sequence.

Note that the exponent of the polynomial factor is an absolute constant (not depending on \(d\) nor on \(\varepsilon\)).

We then apply our framework to \textsc{Coloring} and \textsc{Max Induced Matching}.

\[ \textbf{Theorem 4.} \] For every \(\varepsilon > 0\), \textsc{Coloring} and \textsc{Max Induced Matching} admit polynomial-time \(n^\varepsilon\)-approximation algorithms on \(n\)-vertex graphs of bounded twin-width given with an \(O(1)\)-sequence.

The main additional difficulty for \textsc{Coloring} is that one cannot satisfactorily solve/approximate that problem on a modular decomposition by simply coloring its modules and its quotient graph. One needs to tackle a more general problem called \textsc{Set Coloring}. Fortunately this generalization is the fixed point we are looking for: approximating \textsc{Set Coloring} can be done in our framework by mere recursive calls (to itself).

For \textsc{Max Induced Matching}, we face a new kind of obstacle. It can be the case that no decent solution is contained in any color class \(X\) – in the chosen \(d + 1\)-coloring of the red graph \(G/\mathcal{P}_{\lfloor \sqrt{n} \rfloor}\). For instance, it is possible that any such color class \(X\) induces in \(G\) an edgeless graph, while very large induced matchings exist with endpoints in two distinct color classes. We thus need to also find large induced matchings within the black edges and within the red edges of \(G/\mathcal{P}_{\lfloor \sqrt{n} \rfloor}\). This leads to a more intricate strategy intertwining the coloring of bounded-degree graphs (specifically the red graph and the square of its line graph) and recursive calls to induced subgraphs of \(G\), and to special induced subgraphs of the total graph (i.e., made by both the red and black edges) of \(G/\mathcal{P}_{\lfloor \sqrt{n} \rfloor}\). Although this is not necessary, one can observe that the latter graphs are also induced subgraphs of \(G\) itself.
We then explore the limits of our results and framework in terms of amenable problems. We give the following technical generalization to the approximation algorithms for MIS and Max Induced Matching.

▶ **Theorem 5.** For every connected graph $H$ and $\varepsilon > 0$, Mutually Induced $H$-Packing admits a polynomial-time $n^\varepsilon$-approximation algorithms on $n$-vertex graphs of bounded twin-width given with an $O(1)$-sequence.

In this problem, one seeks for a largest induced subgraph that consists of a disjoint union of copies of $H$. All the previous technical issues are here combined. We try all the possibilities of batching the vertices of $H$ into at most $|V(H)|$ parts of $G/{\mathcal{P}_{\lfloor\sqrt{n}\rfloor}}$, based on the trigraph that these parts define. For instance with $H = K_2$ (an edge), i.e., the case of Max Induced Matching, the three possible trigraphs are the 1-vertex trigraph, two vertices linked by a red edge, and two vertices linked by a black edge. In the general case, the problem generalization is quite delicate to find. We have to keep some partitions of $V(G)$ and $V(H)$ to enforce that the copies of $H$ in $G$ follow a pattern that the algorithm committed to higher up in the recursion tree, and a weight function on $|V(H)|$-tuples of vertices of $G$, not to forget how many mutually induced copies of $H$ can be packed within these vertices. The other novelty is that some recursive calls are on induced subgraphs of the total graph of $G/{\mathcal{P}_{\lfloor\sqrt{n}\rfloor}}$ that are not induced subgraphs of $G$. Fortunately, these graphs keep the same bound of versatile twin-width, and thus our framework allows it.

Defining, for a family of graphs $\mathcal{H}$, Mutually Induced $\mathcal{H}$-Packing as the same problem where the connected components of the induced subgraph should all be in $\mathcal{H}$, we get a similar approximation factor when $\mathcal{H}$ is a finite set of connected graphs. (Note that Mutually Induced $H$-Packing is sometimes called Independent Induced $H$-Packing.) In particular, we can similarly approximate Independent $H$-Packing, which is the same problem but the copies of $H$ need not be induced. (Our approximation algorithms could extend to other $H$-packing variants without the independence requirement, but these problems can straightforwardly be $O(1)$-approximated in general graphs.)

We can handle some cases when $\mathcal{H}$ is infinite, too. For instance, by slightly adapting the case of MIS, we can get an $n^\varepsilon$-approximation when $\mathcal{H}$ is the set of all cliques. We show this more involved example, also expressible as Mutually Induced $\mathcal{H}$-Packing for $\mathcal{H}$ the set of all trees or the set all stars.

▶ **Theorem 6.** For every $\varepsilon > 0$, finding the induced (star) forest with the most edges admits a polynomial-time $n^\varepsilon$-approximation algorithms on $n$-vertex graphs of bounded twin-width given with an $O(1)$-sequence.

As we already mentioned, our framework is exclusively useful for problems that are very inapproximable in general graphs; at least for which an $n^\varepsilon$-approximation algorithm is not known for every $\varepsilon > 0$. Are there natural such problems that cannot be approximated better in graphs of bounded twin-width? We answer this question positively with the example of Min Independent Dominating Set.

▶ **Theorem 7.** For every $\varepsilon > 0$, Min Independent Dominating Set does not admit an $n^{1-\varepsilon}$-approximation algorithm in $n$-vertex graphs given with an $O(1)$-sequence, unless $P=NP$.

The reduction is the same as the one for general graphs [22], but performed from a planar variant of 3-SAT. The obtained instances are not planar but can be contracted to planar trigraphs, hence overall have bounded twin-width.

Finally the case of Longest Path and Longest Induced Path is interesting. The best approximation factor for the former [18] is worse than $n^{0.99}$, while the latter is known
to have the same inapproximability as MIS [31]. However an $n^\varepsilon$-approximation algorithm (for every $\varepsilon > 0$) is not excluded for LONGEST PATH. We show that the property of bounded twin-width is unlikely to help for these two problems, as it would lead to better approximation algorithms for LONGEST PATH in general graphs. This is mainly because subdividing at least $2 \log n$ times every edge of any $n$-vertex graph gives a graph with twin-width at most 4 [3].

**Theorem 8.** For any $r = \omega(1)$, an $r$-approximation for LONGEST INDUCED PATH or LONGEST PATH on graphs given with an $O(1)$-sequence would imply a $(1 + o(1))r$-approximation for LONGEST PATH in general graphs.

In turn, this can be used to exhibit a family $\mathcal{H}$ with an infinite antichain for the induced subgraph relation such that MUTUALLY INDUCED $\mathcal{H}$-PACKING is hard to $n^\varepsilon$-approximate on graphs of bounded twin-width. The family $\mathcal{H}$ is simply the set of all paths terminated by triangles at both ends.

**Theorem 9.** There is an infinite family $\mathcal{H}$ of connected graphs such that if for every $\varepsilon > 0$, MUTUALLY INDUCED $\mathcal{H}$-PACKING admits an $n^\varepsilon$-approximation algorithm on $n$-vertex graphs given with an $O(1)$-sequence, then so does LONGEST PATH on general graphs.

Table 1 summarizes our results and hints at future work.

| Problem name                  | lower bound general graphs | upper bound bounded tww | lower bound bounded tww |
|-------------------------------|----------------------------|-------------------------|-------------------------|
| MAX INDEPENDENT SET           | $n^{1-\varepsilon}$       | $n^\varepsilon$         | ? , self-improvement    |
| COLORING                      | $n^{1-\varepsilon}$       | $n^\varepsilon$         | $4/3 - \varepsilon$     |
| MAX INDUCED MATCHING          | $n^{1-\varepsilon}$       | $n^\varepsilon$         | ?                       |
| MUT. IND. $\mathcal{H}$-PACKING| $n^{1-\varepsilon}$       | $n^\varepsilon$ ($H$ connected) | ?                       |
| MUT. IND. $\mathcal{H}$-PACKING| $n^{1-\varepsilon}$       | $n^\varepsilon$ for some $\mathcal{H}$ | LONGEST PATH-hard |
| MIN IND. DOM. SET             | $n^{1-\varepsilon}$       | $n/\text{polylog}(n)$   | $n^{1-\varepsilon}$    |
| LONGEST PATH                  | $2^{\log^{1-\varepsilon} n}$ | $n/\exp(O(\sqrt{\log n})$ | LONGEST PATH-hard      |
| LONGEST INDUCED PATH          | $n^{1-\varepsilon}$       | $n/\text{polylog}(n)$   | LONGEST PATH-hard      |
| MIN DOMINATING SET            | $(1 - \varepsilon) \ln n$ | $O(1)$                  | ?                       |

For the main highly inapproximable graph problems, we either obtain an $n^\varepsilon$-approximation algorithm on graphs of bounded twin-width given with an $O(1)$-sequence, or a conditional obstruction to such an algorithm. In the former case, can we improve further the approximation factor? The next theorem was observed using the self-improvement reduction of Feige et al. [17], which preserves the twin-width bound. This reduction consists of going from a graph $G$ to the lexicographic product $G'[G]$, where every vertex of $G$ is replaced by a module inducing a copy of $G$ (and iterating this trick).

**Theorem 10 ([7]).** Let $r : \mathbb{N} \to \mathbb{R}$ be any non-decreasing function such that for every $\varepsilon > 0$, $r(n) = o(n^\varepsilon)$. If MAX INDEPENDENT SET admits an $r(n)$-approximation algorithm
on $n$-vertex graphs of bounded twin-width given with an $O(1)$-sequence, then it further admits an $r(n)^2$-approximation.

To our knowledge, the application of the self-improvement trick is always to strengthen a lower bound, and never to effortlessly obtain a better approximation factor. Therefore, we may take Theorem 10 as a weak indication that our approximation ratio is best possible. Still, not even a polynomial-time approximation scheme (PTAS) is ruled out for MAX INDUCED MATCHING, MIN DOMINATING SET, etc.) and we would like to see better approximation algorithms. For COLORING, as was previously observed [7], a PTAS is ruled out by the NP-hardness of deciding if a planar graph is 3-colorable or 4-chromatic, since planar graphs have twin-width at most 9 and a 9-sequence can be found in linear time [24].

2 Preliminaries

For $i$ and $j$ two integers, we denote by $[i,j]$ the set of integers that are at least $i$ and at most $j$. For every integer $i$, $[i]$ is a shorthand for $[1,i]$.

2.1 Handled graph problems

We will consider several problems throughout the paper. We recall here the definition of the most central ones. Some technical problem generalizations will be defined along the way.

**Weighted Max Independent Set (WMIS, for short)**

**Input:** A graph $G$ and a weight function $V(G) \to \mathbb{Q}$.

**Output:** A set $S \subseteq V(G)$ such that $\forall u,v \in S, uv \notin E(G)$ maximizing $w(S) := \sum_{v \in S} w(v)$.

A feasible solution to WMIS is called an independent set. The MAX INDEPENDENT SET (MIS, for short) problem is the particular case with $w(v) = 1, \forall v \in V(G)$. We may denote by $\alpha(G)$, the independence number, that is the optimum value of WMIS on graph $G$.

**Coloring**

**Input:** A graph $G$.

**Output:** A partition $\mathcal{P}$ of $V(G)$ into independent sets minimizing the cardinality of $\mathcal{P}$.

Equivalently, Coloring can be expressed as finding an integer $k$ and a map $c : V(G) \to [k]$ such that for every $uv \in E(G)$, $c(u) \neq c(v)$, while minimizing $k$.

**Max Induced Matching**

**Input:** A graph $G$, possibly together with a weight function $w : E(G) \to \mathbb{Q}$.

**Output:** A set $S \subseteq E(G)$ such that $\forall uv \neq u'v' \in S$, $\{u,v\} \cap \{u',v'\} = \emptyset$ and $G[\{u,v,u',v'\}]$ has exactly two edges, maximizing $w(S) := \sum_{e \in S} w(e)$.

An induced matching is a pairwise disjoint set of edges (i.e., a matching) with no edge bridging them. We now give a common generalization of WMIS and MAX INDUCED MATCHING.

**Mutually Induced $H$-packing**

**Input:** A graph $G$, possibly together with a weight function $w : V(G) \to \mathbb{Q}$.

**Output:** A set $S \subseteq V(G)$ such that $G[S]$ is a disjoint union of graphs each isomorphic to a graph in $H$, maximizing $w(S) := \sum_{v \in S} w(v)$.

When $H$ consists of a single graph, say $H$, we simply denote the former problem Mutually Induced $H$-packing. WMIS and MAX INDUCED MATCHING are the special cases when $H$ is a vertex and an edge, respectively.
2.2 The contraction and partition viewpoints of twin-width

A trigraph \( G \) has vertex set \( V(G) \), black edge set \( E(G) \), red edge set \( R(G) \) such that
\[ E(G) \cap R(G) = \emptyset \quad \text{and} \quad E(G), R(G) \subseteq \binom{V(G)}{2}. \]
A contraction in a trigraph \( G \) replaces a pair of (non-necessarily adjacent) vertices \( u, v \in V(G) \) by one vertex \( u' \) that is linked to \( G \setminus \{u, v\} \) in the following way to form a new trigraph \( G' \).

For every \( z \in V(G) \setminus \{u, v\} \), \( wz \in E(G') \) whenever \( uz, vz \in E(G) \), \( wz \notin E(G') \cup R(G') \) whenever \( uz, vz \notin E(G) \cup R(G) \), and \( wz \in R(G') \), otherwise. The red graph \( (V(G), R(G)) \) will be denoted by \( R(G) \). We denote by \( T(G) \) the total graph of \( G \) defined as \( (V(G), E(G) \cup R(G)) \).

An induced subtrigraph of a trigraph \( G \) is obtained by removing vertices (but no edges) to \( G \), analogously to induced subgraphs. A partial contraction sequence of an \( n \)-vertex (tri)graph \( G \) (to a trigraph \( H \)) is a sequence of trigraphs \( G = G_n, \ldots, G_1 = H \) for some \( t \in [n] \) such that \( G_i \) is obtained from \( G_{i+1} \) by performing one contraction. A (complete) contraction sequence is such that \( t = 1 \), that is, \( H \) is the 1-vertex trigraph. A \( d \)-sequence \( S \) of \( G \) is a contraction sequence of \( G \) in which the red graph of every trigraph of \( S \) has maximum degree at most \( d \).

Assume that there is a partial contraction sequence from a (tri)graph \( G \) to a trigraph \( H \).

If \( u \) is a vertex of \( H \), then \( u(G) \subseteq V(G) \) denotes the set of vertices eventually contracted into \( u \) in \( H \). We denote by \( P(H) \) the partition \( \{u(G) : u \in V(H)\} \) of \( V(G) \). If \( G \) is clear from the context, we may refer to a part of \( H \) as any set in \( \{u(G) : u \in V(H)\} \). We will mostly see \( d \)-sequences as sequences of partitions, that is, \( P_n, \ldots, P_1 \) with \( P_i := \{u(G) : u \in V(G_i)\} \) when \( G_n, \ldots, G_1 \) is a partial (contraction) \( d \)-sequence.

Given a graph \( G \) and a partition \( P \) of \( V(G) \), the quotient graph of \( G \) with respect to \( P \) is the graph with vertex set \( P \), where \( PP' \) is an edge if there is \( u \in P \) and \( v \in P' \) such that \( uv \in E(G) \). Given a (tri)graph \( G \) and a partition \( P \) of \( V(G) \), the quotient trigraph \( G/P \) is the trigraph with vertex set \( P \), where \( PP' \) is a black edge if these two parts are fully adjacent for every \( u \in P \) and every \( v \in P' \), \( uv \in E(G) \), and a red edge if either there is \( u \in P \) and \( v \in P' \) such that \( uv \in R(G) \), or there is \( u_1, u_2 \in P \) and \( v_1, v_2 \in P' \) such that \( u_1v_1 \in E(G) \) and \( u_2v_2 \notin E(G) \).

A trigraph \( H \) is a cleanup of another trigraph \( G \) if \( V(H) = V(G) \), \( R(H) \subseteq R(G) \), and \( E(G) \subseteq E(H) \subseteq E(G) \cup R(G) \). That is, \( H \) is obtained from \( G \) by turning some of its red edges into black edges or non-edges. We further say that \( H \) is full cleanup of \( G \) if \( H \) has no red edge, and thus, is considered as a graph. Note that the total graph \( T(G) \) and the black graph \( (V(G), E(G)) \) of a trigraph \( G \) are extreme examples of full cleanups of \( G \).

2.3 Balanced partition sequences

The notion of versatile twin-width is a crucial opening step to our algorithms; see [5]. Let us call \( d \)-contraction a contraction between two trigraphs of maximum red degree at most \( d \).

A tree of \( d \)-contractions of a trigraph \( G \) (of maximum red degree at most \( d \)) is a rooted tree, whose root is labeled by \( G \), whose leaves are all labeled by 1-vertex trigraphs \( K_1 \), and such that one can go from any parent to any of its children by performing a single \( d \)-contraction. Observe that \( d \)-sequences coincide with trees of \( d \)-contractions that are paths. A trigraph \( G \) has versatile twin-width \( d \) if \( G \) admits a tree of \( d \)-contractions in which every internal node, labeled by, say, \( F \), has at least \( |V(F)|/d \) children each obtained by contracting one of a list of \( |V(F)|/d \) pairwise disjoint pairs of vertices of \( F \).

It was shown that twin-width and versatile twin-width are functionally equivalent [5].

The relevant consequence for our purposes is that every graph \( G \) with a \( d \)-sequence admits a balanced \( d \)-sequence, where \( d = h(d') \) depends only on \( d' \), i.e., one for which the partitions \( P_n, \ldots, P_1 \) are such that for every \( i \in [n] \) and \( P \in P_i \), \( |P| \leq d \cdot \frac{2}{7} \). As we will resort to
recursion on induced subtrigraphs and quotient trigraphs, we need to keep more information on those substances that the mere fact that they have twin-width at most \(d\) (otherwise the twin-width bound could quickly diverge).

This will be done by opening up the proof in [5], and handling divided 0,1,\(r\)-matrices with some specific properties. Thus we need to recall the relevant definitions.

Given two partitions \(\mathcal{P}, \mathcal{P}'\) of the same set, we say that \(\mathcal{P}'\) is a coarseening of \(\mathcal{P}\) if every part of \(\mathcal{P}\) is contained in a part of \(\mathcal{P}'\), and \(\mathcal{P}, \mathcal{P}'\) are distinct. Given a matrix \(M\), we call row division (resp. column division) a partition of the rows (resp. columns) of \(M\) into parts of consecutive rows (resp. columns). A \((k, \ell)\)-division, or simply division, of a matrix \(M\) is a pair \((\mathcal{R} = \{R_1, \ldots, R_k\}, \mathcal{C} = \{C_1, \ldots, C_\ell\})\) where \(\mathcal{R}\) is a row division and \(\mathcal{C}\) is a column division.

In a matrix division \((\mathcal{R}, \mathcal{C})\), each part \(R \in \mathcal{R}\) is called a row part, and each part \(C \in \mathcal{C}\) is called a column part. Given a subset \(R\) of rows and a subset \(C\) of columns in a matrix \(M\), the zone \(M[R, C]\) denotes the submatrix of all entries of \(M\) at the intersection between a row of \(R\) and a column of \(C\). A zone of a matrix partitioned by \((\mathcal{R}, \mathcal{C}) = ((\mathcal{R}_1, \ldots, \mathcal{R}_k), (\mathcal{C}_1, \ldots, \mathcal{C}_\ell))\) is any \(M[R_i, C_j]\) for \(i \in [k]\) and \(j \in [\ell]\). A zone is constant if all its entries are identical, horizontal if all its columns are equal, and vertical if all its rows are equal. A 0,1-corner is a \(2 \times 2\) 0,1-matrix which is neither horizontal nor vertical.

Unsurprisingly, 0,1,\(r\)-matrices are such that each entry is in \(\{0, 1, r\}\) where \(r\) is an error symbol that should be understood as a red edge. A neat division of a 0,1,\(r\)-matrix is a division for which every zone either contains only \(r\) entries or contains no \(r\) entry and is horizontal or vertical (or both, i.e., constant). Zones filled with \(r\) entries are called mixed. A neatly divided matrix is a pair \((M, (\mathcal{R}, \mathcal{C}))\) where \(M\) is a 0,1,\(r\)-matrix and \((\mathcal{R}, \mathcal{C})\) is a neat division of \(M\). A \(t\)-mixed minor in a neatly divided matrix is a \((t, t)\)-division which coarsens the neat subdivision, and contains in each of its \(t^2\) zones at least one mixed zone (i.e., filled with \(r\) entries) or a 0,1-corner. A neatly divided matrix is said \(t\)-mixed free if it does not admit a \(t\)-mixed minor.

A mixed cut of a row part \(R \in \mathcal{R}\) of a neatly divided matrix \((M, (\mathcal{R}, \mathcal{C} = \{C_1, C_2, \ldots\}))\) is an index \(i\) such that both \(M[R, C_i]\) and \(M[R, C_{i+1}]\) are not mixed, and there is a 0,1-corner in the 2-by-\(|R|\) zone defined by the last column of \(C_i\), the first column of \(C_{i+1}\), and \(R\). The mixed value of a row part \(R \in \mathcal{R}\) of a neatly divided matrix \((M, (\mathcal{R}, \mathcal{C} = \{C_1, C_2, \ldots\}))\) is the number of mixed zones \(M[R, C_i]\) plus the number of mixed cuts between two (adjacent non-mixed) zones \(M[R, C_i]\) and \(M[R, C_{i+1}]\). We similarly define the mixed value of a column part \(C \in \mathcal{C}\). The mixed value of a neat division of a 0,1,\(r\)-matrix is the maximum of the mixed values taken over every part. The part size of a division \((\mathcal{R}, \mathcal{C})\) is defined as \(\max(\max_{R \in \mathcal{R}} |R|, \max_{C \in \mathcal{C}} |C|)\). A division is symmetric if the largest row index of each row part and the largest column index of each column part define the same set of integers. We call symmetric fusion of a symmetric division the fusion of two consecutive parts in \(\mathcal{C}\) and of the two corresponding parts in \(\mathcal{R}\). A symmetric fusion on a symmetric division yields another symmetric division. A matrix \(A := (a_{i,j})_{i,j}\) is said symmetric in the usual sense, namely, for every entry \(a_{i,j}\) of \(A\), \(a_{i,j} = a_{j,i}\).

In what follows, we set \(c_d := 8/3(t + 1)^22^{4t}\). The following definition is key.

\begin{definition}
Let \(\mathcal{M}_{n,d}\) be the class of the neatly divided \(n \times n\) symmetric 0,1,\(r\)-matrices \((M, (\mathcal{R}, \mathcal{C}))\), such that \((\mathcal{R}, \mathcal{C})\) is symmetric and has:
\begin{itemize}
  \item mixed value at most \(4c_d\),
  \item part size at most \(2^{4c_d + 2}\), and
  \item no \(d\)-mixed minor.
\end{itemize}

The red number of a matrix is the maximum number of \(r\) entries in a single column or row of the matrix.
\end{definition}
\textbf{Lemma 12.} Let \((M, (R, C)) \in \mathcal{M}_{n,d}\). The red number of \(M\) is at most \(c_d \cdot 2^{4c_d+4}\). Thus, the trigraph whose adjacency matrix is \(M\) has maximum red degree at most \(c_d \cdot 2^{4c_d+4}\).

\textbf{Proof.} Any row or column intersects at most \(4c_d\) mixed zones (filled with \(r\) entries). Each mixed zone has width and length bounded by the part size \(2^{4c_d+2}\). Hence the maximum total number of \(r\) entries on a single row or column is at most \(4c_d \cdot 2^{4c_d+2} = c_d \cdot 2^{4c_d+4}\).

A coarsening of a neatly divided matrix \((M, (R, C))\) is a neatly divided matrix \((M', (R', C'))\) such that \((R', C')\) is a coarsening of \((R, C)\), and \(M'\) is obtained from \(M\) by setting to \(r\) all entries that lie, in \(M\) divided by \((R', C')\), in a zone with at least one \(r\) entry or a 0,1-corner.

We also refer to the process of going from \((M, (R, C))\) to \((M', (R', C'))\) as coarsening operation. A coarsening operation from \((M, (R, C)) \in \mathcal{M}_{n,d}\) to \((M', (R', C'))\) is said invariant-preserving if \((M', (R', C')) \in \mathcal{M}_{n,d}\).

The following lemma is the crucial building block of the current section.

\textbf{Lemma 13 ([6, Lemma 18]).} We set \(s := 2^{4c_d+4}\). Every neatly divided matrix \((M, (R, C)) \in \mathcal{M}_{n,d}\) has an invariant-preserving coarsening \((M', (R', C')) \in \mathcal{M}_{n,d}\) with \([n/s]\) disjoint pairs of identical columns. Given \((M, (R, C))\), both \((M', (R', C'))\) and the pairs of columns can be computed in \(n^{O(1)}\) time.

In [6], it is not explicitly stated that the invariant-preserving coarsening (hence the pairs of identical columns) can be found in polynomial time. However it is easy to check that the proof is effective, since it greedily symmetrically fuses two consecutive parts, provided the resulting divided matrix remains in \(\mathcal{M}_{n,d}\). A special case of the following observation is shown in [6, Lemma 19].

\textbf{Lemma 14.} Let \((M, (R, C)) \in \mathcal{M}_{n,d}\) be a neatly divided matrix. Removing a set of \(h\) columns and the \(h\) corresponding rows, and possibly removing from the division the parts that are now empty, results in a neatly divided matrix in \(\mathcal{M}_{n-h,d}\).

\textbf{Proof.} By construction, the new matrix and division are symmetric. The new neatly divided matrix remains \(d\)-mixed free. The part size and the mixed value can only decrease.

\textbf{Lemma 15 ([6, Beginning of Lemma 20]).} Given any graph \(G\) with a \(d\)-sequence, one can find in polynomial-time an adjacency matrix \(M\) of \(G\), such that \((M, (R, C))\) is a neatly divided matrix of \(\mathcal{M}_{n,2d+2}\) with \((R, C)\) the finest division of \(M\) (i.e., the one where all parts are of size 1).

The adjacency matrix of a trigraph extends the one of a graph by putting \(r\) symbols when the vertices of the corresponding row and column are linked by a red edge. A neatly divided matrix \((M, (R, C))\) is said conform to a trigraph \(G\) if \(M\) is the adjacency matrix of a trigraph \(G'\) such that \(G\) is a cleanup of \(G'\). Furthermore, we assume (and keep implicit) that we know the one-to-one correspondence between each row (and corresponding column) of \(M\) and vertex of \(G\).

\textbf{Lemma 16.} Let \(d\) be a natural, \(s := 2^{4c_d+4}\), and \(d' := c_d \cdot 2^{4c_d+4}\). Let \(G\) be an \(n\)-vertex trigraph given with a neatly divided matrix \((M, (R, C)) \in \mathcal{M}_{n,d}\) conform to \(G\). A partial \(d'\)-sequence \(S\) from \(G\) to a trigraph \(H\) satisfying

\[ |V(H)| = \lfloor \sqrt{n} \rfloor, \quad \text{and} \quad \forall u \in V(H), |u(G)| \leq s \sqrt{n}, \]

and a neatly divided matrix \((M', (R', C')) \in \mathcal{M}_{\lfloor \sqrt{n} \rfloor, d'}\) conform to \(H\) can be computed in time \(n^{O(1)}\).
Given a neatly divided matrix \( M \) and a partition \( \mathcal{P} = \{P_1, P_2, \ldots, P_{\sqrt{n}}\} \) of \( V(G) \) satisfying that, for every integer \( 1 \leq i \leq \lfloor \sqrt{n} \rfloor \), \( |P_i| \leq s \sqrt{n} \leq d' \sqrt{n} \), and

the red graph of \( G/\mathcal{P} \) has maximum degree at most \( d' \).

We will need a stronger inductive form of Lemma 17, also a consequence of Lemmas 15 and 16.

\[ \text{Lemma 18.} \ \text{Let} \ \tilde{d} \ \text{be a natural,} \ \tilde{d} = 2\tilde{d} + 2, \ \text{and set} \ s := 2^{4s\tilde{d}+4}. \ \text{Given an} \ n \text{-vertex graph} \ G \ \text{with a} \ \tilde{d} \text{-sequence, one can compute in time} \ n^{O(1)} \ \text{a partition} \ \mathcal{P} = \{P_1, P_2, \ldots, P_{\lfloor \sqrt{n} \rfloor}\} \ \text{of} \ V(G) \ \text{satisfying that, for every integer} \ 1 \leq i \leq \lfloor \sqrt{n} \rfloor, \ |P_i| \leq s \sqrt{n} \leq d' \sqrt{n}, \ \text{and} \]

\[ \text{the red graph of} \ G/\mathcal{P} \ \text{has maximum degree at most} \ d'. \]

Proof. If we are given a graph \( G \) with a \( \tilde{d} \)-sequence, we immediately compute a neatly divided matrix \( (M, (R, C)) \in \mathcal{M}_{n,d} \) conform to \( G \), by Lemma 15. We then proceed as if we received the second kind of input.

We will build iteratively the partition \( \mathcal{P} = \{P_1, P_2, \ldots, P_{\lfloor \sqrt{n} \rfloor}\} \) starting from the finest partition. At each step we merge two parts, until the number of parts is \( \lfloor \sqrt{n} \rfloor \). At this point, we have the desired partition \( \mathcal{P} \).

We iteratively maintain a trigraph \( G^z \) and a neatly divided matrix \( (M^z, (R^z, C^z)) \in \mathcal{M}_{n-2z+1,d} \) conform to it. The maintained partition is just the one corresponding to the parts of \( G^z \). Initially, \( G^z \) is \( G \), and \( (M^1, (R^1, C^1)) = (M, (R, C)) \in \mathcal{M}_{n,d} \). At step \( z \) we do the following. We apply Lemma 13 on \( (M^z, (R^z, C^z)) \in \mathcal{M}_{n-2z+1,d} \) and obtain, in polynomial-time, an invariant-preserving coarsening \( (M^{z+1}, (R^{z+1}, C^{z+1})) \in \mathcal{M}_{n-2z+1,d} \) and an induced subtrigraph \( (M^{z+1}, (R^{z+1}, C^{z+1})) \in \mathcal{M}_{\lfloor \sqrt{n} \rfloor, d} \) conform to \( H \) can be computed in time \( n^{O(1)} \).

Proof. This is a consequence of Lemmas 13 and 14; see the proof of the more general Lemma 18.
an induced subgraph $H'$ of $G/P$ (i.e., removing vertices from it), we get, by removing the corresponding rows and columns in $(M'(\sqrt{n}), (R'(\sqrt{n}), C'(\sqrt{n})))$ a neatly divided matrix $(M'(\sqrt{n}), (R'(\sqrt{n}), C'(\sqrt{n}))) \in M_{\lfloor \sqrt{n} \rfloor}$ conform to $H'$, by Lemma 14. Note finally that taking a cleanup $H$ of $H'$, we can simply keep $(M'(\sqrt{n}), (R'(\sqrt{n}), C'(\sqrt{n})))$ as a neatly divided matrix of $M_{\lfloor \sqrt{n} \rfloor}$ conform to $G$. The second item, concerning induced subgraphs $G[\bigcup_{i \in J} P_i]$ is a simple application of Lemma 14, and works more generally for any induced subgraph of $G$. ▶

In effect, we will only apply Lemma 18 for graphs $G$ and $H$, i.e., when $H$ is an induced subgraph of $G$ or a full cleanup of an induced subgraph of $G/P$. Indeed, the structures $H$ will correspond to subinstances. We want those to be graphs, so that the tackled graph problem is well-defined on them.

## 3 Approximation algorithms for Max Independent Set

We naturally start our study with Max INDEPENDENT SET, a central problem that is very inapproximable [23, 34], and yet constitutes the textbook example of our approach.

### 3.1 Subexponential-time constant-approximation algorithm

We present a subexponential-time $O_d(1)$-approximation for WMIS on graphs given with a $d$-sequence, which we recall, is unlikely to exist in general graphs [14].

**Lemma 19.** Let $d'$ be a natural, $s := 2^{4d'+4}$, and $d := c_d \cdot 2^{4d'+4}$. Assume $n$-vertex inputs $G$, vertex-weighted by $w$, are given with a $d'$-sequence. WEIGHTED MAX INDEPENDENT SET can be $(d+1)$-approximated in time $2^{O_d(\sqrt{n})}$ on these inputs.

**Proof.** By Lemma 17, we compute in polynomial time a partition $P = \{P_1, \ldots, P_{\lfloor \sqrt{n} \rfloor} \}$ of $V(G)$ whose parts have size at most $s \sqrt{n}$ and such that $R(G/P)$ has maximum degree at most $d$.

For every integer $1 \leq i \leq \lfloor \sqrt{n} \rfloor$, we compute a heaviest independent set in $G[P_i]$, say $S_i$. Even with an exhaustive algorithm, this takes time $\sqrt{n} \cdot s^2 \sqrt{n} \cdot 2^{4 \sqrt{n}} = 2^{O_d(\sqrt{n})}$. We then $(d+1)$-color (in linear time) $R(G/P)$, which is possible since this graph has maximum degree at most $d$. This defines a coarsening of $P$ in $d+1$ parts $Q = \{C_1, \ldots, C_{d+1} \}$. Thus, $Q$ is a partition of $V(G)$ such that $C_j$ consists of all the parts $P_i \in P$ receiving color $j$ in the $(d+1)$-coloring of $R(G/P)$.

For every $j \in [d+1]$, let $H_j$ be the graph $(G/P)[C_j]^4$ vertex-weighted by $P_i \subseteq C_j \mapsto w(S_i)$. Note that $(G/P)[C_j]$ can indeed be assimilated to a graph, since it has, by design, no red edge. We compute a heaviest independent set in $H_j$, say $R_j$. This takes time $(d+1) n \cdot 2^{\sqrt{n}} = 2^{O_d(\sqrt{n})}$. We output $\bigcup_{P_i \subseteq R_j} S_i$ for the index $j \in [d+1]$ maximizing $\sum_{P_i \subseteq R_j} w(S_i)$.

This finishes the description of the algorithm. We already argued that its running time is $2^{O_d(\sqrt{n})}$. We shall justify that it does output an independent set of weight at least a $\frac{1}{3^{d+1}}$ fraction of the optimum $\alpha(G)$.

* $I$ is indeed an independent set. For any $j \in [d+1]$, consider two vertices $x, y \in \bigcup_{P_i \subseteq R_j} S_i$. If $\{x, y\} \in S_i$ for some $i$, then $x$ and $y$ are non-adjacent since $S_i$ is an independent set of $G[P_i]$. Else $x \in S_i$ and $y \in S_{i'}$ for some $i \neq i'$. $P_i$ and $P_{i'}$ are not linked by a black edge in $(G/P)[C_j]$ since $R_j$ is an independent set in $H_j$, nor they can be linked by a red edge (there are none in $(G/P)[C_j]$). Thus again, $x$ and $y$ are non-adjacent in $G$.

---

4 We use this notation as a slight abuse of notation for $(G/P)[\{P_i : P_i \subseteq C_j \}]$.
Figure 2 The trigraph $G/P$ with its $\lfloor \sqrt{n} \rfloor$ vertices, each corresponding to a subset of at most $s\sqrt{n}$ vertices of $G$. The weights $w(S)$ of heaviest independent sets $S_i$ of $G[P_i]$ for each part $P_i$ of the color class $C_2$ of the $d+1$-coloring of $R(G/P)$. A heaviest independent set in the so-weighted $(G/P)[C_2]$ (shaded) corresponds to an optimum solution in $G[\bigcup P_i \subseteq C_2]$. One of these $d+1$ independent sets is a $d+1$-approximation.

3.2 Improving the approximation factor

We notice in this short section that the approximation factor of Lemma 19 can be improved using the notion of clustered coloring. The clustered chromatic number of a class of graphs is the smallest integer $k$ such that there is a constant $c$ for which all the graphs of the class can be $k$-colored such that every color class induces a subgraph whose connected components have size at most $c$. A proper coloring is a particular case of clustered coloring when $c = 1$.

Instead of properly coloring the red graph, as we did in the proof of Lemma 19, we could use less colors and allow for small monochromatic components (in place of monochromatic components of size 1). We use for that the following bound due to Alon et al.

\begin{theorem}[\cite{1}] The class of graphs of maximum degree at most $d$ has clustered chromatic number at most $\lceil \frac{d+2}{3} \rceil$.
\end{theorem}

We can use this lemma to improve our approximation algorithms.

\begin{theorem} On inputs as in Lemma 19 with $s := 2^{4c_d + 4}$, and $d := c_d \cdot 2^{4c_d + 4}$, \textsc{Weighted Max Independent Set} further admits an $\lceil \frac{d+2}{3} \rceil$-approximation algorithm in time $2^{O_\epsilon(\sqrt{n})}$.
\end{theorem}

Proof. Again, we compute in polynomial time a partition $\mathcal{P} = \{P_1, \ldots, P_{\lfloor \sqrt{n} \rfloor}\}$ of $V(G)$ whose parts have size at most $s\sqrt{n}$ and such that $R(G/\mathcal{P})$ has maximum degree at most $d$,
using Lemma 17. Let $c$ be the constant such that $\mathcal{R}(G/P)$ admits a clustered coloring using $\left\lfloor \frac{\log n}{d+2} \right\rfloor$ colors such that each color class $C_j$ (with $j \in \left[\frac{\log n}{d+2}\right]$) is such that the connected components $C^h_j, C^h_{j+1}, \ldots, C^h_{j+2}$ of $\mathcal{R}(G/P)[C_j]$ have size at most $c \cdot d$ each. This coloring is guaranteed to exist by Theorem 20. Due to the overall running time, we might as well compute it by exhaustive search, in time $2^{O_d(\sqrt n)}$.

For every $j, h \in \left[\frac{\log n}{d+2}\right]$ and $i \in [d+2]$, we denote the connected component $C^h_i \subseteq C_j$ of $\mathcal{R}(G/P)[C_j]$ by $P_{i}$. For every $i \in P$, that are included in $C^h_j$. For every $j \in \left[\frac{\log n}{d+2}\right]$, every $i \in [d+2]$, and every $J \subseteq [d+2]$, we compute a heaviest independent set in $\bigcup_{i \in P} P_j(C^h_i)$, which we denote by $S_{j, h, J}$. This takes time $O(\sqrt n \cdot 2^c \cdot 2^{q_3 \sqrt n})$ since $\bigcup_{i \in P} P_j(C^h_i) \leq c \cdot s \cdot \sqrt n$.

For each $C_j$, in time $(2^c)^{\sqrt n} = 2^{c \sqrt n}$, we exhaustively try all subsets $X \subseteq \bigcup_{P \in C_j} P_j$ that are unions of $S_{j, h, J}$ filtering them out when $G[X]$ is not edgeless, and keep a heaviest of them, say $R_j$. Since there can be only be black edges or non-edges between some $P_i \in C^h_j$ and $P'_i \in C^h_j$ with $h \neq h'$, it is clear that a heaviest independent set of $\bigcup_{P \in C_j} P_j$ is indeed a union of $S_{j, h, J}$ (with fixed $j$). We output a heaviest set among the $R_j$s, which is the desired $\left\lfloor \frac{\log n}{d+2}\right\rfloor$-approximation. The running time is as claimed.

### 3.3 Time-approximation trade-offs

Lemma 19 and Theorem 21 run exhaustive algorithms on induced subgraphs of size $O_d(\sqrt n)$. As such, the latter inputs keep the same twin-width upper bound. To speed up the algorithm (admittedly while worsening the approximation factor) it is tempting to recursively call our very algorithm. We show that this leads to a time-approximation trade-off parameterized by an integer $q = 0, \ldots, O_d(\log \log n)$. At one end of this discrete curve, one finds the exact exponential algorithm ($q = 0$), and more interestingly the $d+1$-approximation in time $2^{O_d(\sqrt n)} (q = 1)$, while at the other end lies a polynomial-time algorithm with approximation factor $n^\varepsilon$, where $\varepsilon > 0$ can be made as small as desired.

As we will deal with the same kind of recursions for several problems, we show the following generic abstraction.

#### Lemma 22. Let $d$ be a natural, $d' = 2d + 2$, and $d := cd' \cdot 2^{4c_d+4}$. Let $\Pi$ be an optimization graph problem where inputs come with a $d$-sequence of their $n$-vertex graph $G$, or with a neatly divided matrix $(M, (\mathcal{R}, \mathcal{C})) \in \mathcal{M}_{n,d'}$ conform to $G$. Let $P$ be the partition of $V(G)$ given by Lemma 18. Assume that

1. $\Pi$ can be exactly solved in time $2^{O(n)}$, and there are constants $c_1, c_2, c_3$, and a function $f \geq 1$ such that
2. $d'^3 \cdot 2^2$-approximation of $\Pi$ on $G$ can be built in time $n^{c_2}$ by using at most $n^{c_1}$ calls to an $r$-approximation of $\Pi$ — or another optimization problem $\Pi'$ already satisfying the conclusion of the lemma — on an induced subgraph of $G$ with at most $f(d') \sqrt n$ vertices or a full cleanup of an induced subtrigraph of $G/P$ (on at most $\sqrt n$ vertices).

Then $\Pi$ can be $d'^3(2^{2d-1})$-approximated in time

$$f(d')^{q_1(n)^{(2-2^{-q})(c_1+c_2)}} \cdot 2f(d')^{2(1-2^{-q})}n^{2^{-q}}$$

for any non-negative integer $q$.

#### Proof. The proof is by induction on $q$. The case $q = 0$ is implied by Item 1. The case $q = 1$, and the induction step in general, is nothing more than an abstraction of Lemma 19, where exhaustive algorithms are replaced by recursive calls.

For any $q \geq 0$, we assume that $\Pi$ can $d'^3(2^{2d-1})$-approximated in the claimed running time, and show the same statement for the value $q + 1$. Following Item 2, we run this algorithm — or
one for another optimization problem \( \Pi' \) satisfying the conclusion of the lemma— at most \( n^{c_1} \) times on \( f(d)\sqrt{n} \)-vertex induced subgraphs of the input graph \( G \) or on full cleanups of induced subgraphs of \( G/\mathcal{P} \). The latter graphs have at most \( \sqrt{n} \leq f(d)\sqrt{n} \) vertices. By Lemma 18, we can compute in polynomial time a neatly divided matrix \( (M', (\mathcal{R}', \mathcal{C}')) \in \mathcal{M}[V(H)], d' \) conform to \( H \), for each graph \( H \) of a recursive call; hence the induction applies.

Overall this takes time at most

\[
n^{c_1} + n^{c_2} \cdot \left( (f(d)^q, f(d)\sqrt{n}(2-2^{-q})(c_1+c_2), 2(f(d)\sqrt{n})^{(1-2^{-q})}(f(d)\sqrt{n})^{2^{-q}}) \right)
\]

\[
\leq (f(d)^q n)^{c_1+c_2+\frac{1}{2}(2-2^{-q})(c_1+c_2)} \cdot 2(f(d)^{1-2^{-q}}+2^{-q} n^{\frac{2-2^{-q}}{2}})
\]

\[
= (f(d)^{q+1} n)^{(2-2^{-q})(c_1+c_2)} \cdot 2f(d)^{2-2^{-q}+2^{-q} n^{2-(q+1)}}
\]

\[
= (f(d)^{q+1} n)^{(2-2^{-q+1})(c_1+c_2)} \cdot 2f(d)^{2-2^{-q-1} n^{2-(q+1)}}
\]

For the first inequality, we assume that the two summands are larger than 2, so their sum can be bounded by their product.

Besides we get an approximation of factor at most \((d^q(2^q-1))^2 d^{q+1} = d^{q+1} \). ▶

In more legible terms we have proved that:

\[\textbf{Lemma 23.} \text{Problems } \Pi \text{ satisfying the assumptions of Lemma 22 can be } d^{O(1)(2^q-1)} \text{ approximated in time } 2^{O(d,q) (\sqrt{n})}. \text{ for any non-negative integer } q.\]

If most graph problems admit single-exponential algorithms, we will deal with such a problem that is only known to be solvable in time \( 2^{O(n \log n)} \). Therefore we prove a variant of Lemma 22 with a slightly worse running time.

\[\textbf{Lemma 24.} \text{Let } \Pi \text{ be solvable in time } 2^{O(n \log n)} \text{ and satisfy the second item of Lemma 22. Then } \Pi \text{ can be } d^{O((2^q-1))}-\text{approximated in time}
\]

\[2^{(c_1+e_2)(2-2^{-q}) \log f(d)+f(d)(2-2^{-q}) n^{2-q}} \log n,
\]

for any non-negative integer \( q \).

\[\text{Proof.} \text{ We follow the proof of Lemma 22 when the induction now gives a running time of}
\]

\[n^{c_2} + n^{c_1} \cdot 2^{(c_1+e_2)(2-2^{-q}) \log f(d)+f(d)(2-2^{-q}) n^{2-q}} \log n
\]

\[\leq 2^{(c_1+e_2)(2-2^{-q+1}) \log f(d)+f(d)(2-2^{-q+1}) n^{2-(q+1)}} \log n.
\]

Again the previous lemma can be rewritten as:

\[\textbf{Lemma 25.} \text{Problems } \Pi \text{ satisfying the assumptions of Lemma 24 can be } d^{O(1)(2^q-1)} \text{ approximated in time } 2^{O(d,q) (\sqrt{n} \log n)}. \text{ for any non-negative integer } q.\]

We derive from Lemma 24 the following notable regimes.

\[\textbf{Theorem 26.} \text{Problems } \Pi \text{ satisfying the assumptions of Lemma 24 admit polynomial-time } n^{\varepsilon}-\text{approximation algorithms, for any } \varepsilon > 0.\]
Proof. This is the particular case \( q = \lceil \log \frac{\epsilon \log n}{c_3 \log d} \rceil \).

Indeed the approximation factor is then at most \( d^{c_3(2^\epsilon-1)} \leq d^{c_3 \frac{\epsilon \log n}{c_3 \log d}} = 2^\epsilon \log n \),
while the running time is at most
\[
2^{(c_1+c_2)(2-2^{-\epsilon}) \log f(d)+f(d)(1-2^{-\epsilon}) n^{2-\epsilon}} \log n
\leq 2^{2(c_1+c_2) \log f(d)+f(d) n^{2-\epsilon}} \log n
\]
\[
= n^{2(c_1+c_2) \log f(d)+f(d)n^{2-\epsilon}}.
\]

If further \( II \) can be solved exactly in time \( 2^{O(n)} \) (hence satisfies the assumptions of Lemma 22), one obtains a better running time, where the exponent of \( n \) does not depend on \( \epsilon \). Indeed,
\[
(f(d)^n)^{(2-2^{-\epsilon})(c_1+c_2)} f(d)^{2-2^{-\epsilon}} n^{2-\epsilon} \leq \left( \frac{\epsilon \log n}{c_3 \log d} \right)^{2(c_1+c_2) \log f(d)} f(d)n^{2-\epsilon} \]
\[
\leq 2^{2(c_1+c_2) \log f(d)+f(d)n^{2-\epsilon}} \log n.
\]

\[\blacktriangleleft\]

**Theorem 27.** Problems \( II \) satisfying the assumptions of Lemma 22, resp. Lemma 24, admit a \( \log n \)-approximation algorithm running in time \( 2^{O_d(n \frac{1}{m+d+\epsilon})} \), resp. \( 2^{O_d(n \frac{1}{m+d+\epsilon}) \log n} \).

**Proof.** This is the particular case \( q = \lceil \log \left( \frac{\log \log n}{c_3 \log d} + 1 \right) \rceil \).

This value is computed such that the approximation factor \( d^{\epsilon(2^\epsilon-1)} \) is at most \( \log n \). It can be easily checked that the running times are as announced. \[\blacktriangleleft\]

We derive the following for **Weighted Max Independent Set**.

**Theorem 28.** **Weighted Max Independent Set** on \( n \)-vertex graphs \( G \) (vertex-weighted by \( w \)) given with a \( d \)-sequence satisfies the assumptions of Lemma 22. In particular, this problem admits

- a \( (d+1)^{2^\epsilon-1} \)-approximation in time \( 2^{O_d(n \frac{1}{m+d+\epsilon})} \), for every integer \( q \geq 0 \),
- an \( n^{\epsilon^2} \)-approximation in polynomial-time \( O_d(1) \log O_d(1) n \cdot n^{O(1)} \), for any \( \epsilon > 0 \), and
- a \( \log n \)-approximation in time \( 2^{O_d(n \frac{1}{m+d+\epsilon})} \),

with \( d := c_2 d' + 2 \cdot 2^{4c_2 d' + 2^{-\epsilon}} \).

**Proof.** Even the exhaustive algorithm exactly solves WMIS in time \( 2^{O(n)} \). We thus focus on showing that WMIS satisfies the second item of Lemma 22. We set \( c_1 \geq 1 \) as the required exponent to turn a \( d \)-sequence into a neatly divided matrix of \( M_{n,2d'+2} \) conform to \( G \),

\[ c_2 = \frac{1}{2} + \eta \] for any fixed \( \eta > 0 \), the appropriate \( 1 < c_3 \leq 2 \), and \( f(d) = d \geq 1 \).

The algorithm witnessing the second item is simply the proof of Lemma 19. We first check that this algorithm makes \( \lceil \sqrt{n} \rceil + d + 1 \) recursive calls on induced subgraphs of the input \( G \) each of the \( \lceil \sqrt{n} \rceil \) graphs \( G[P_i] \) where \( P_i \) has indeed size at most \( O_d(\sqrt{n}) \), and each of the \( d+1 \) graphs \( (G/P)[C_j] \) (indeed an induced subgraph of \( G \) by definition of the black graph of a trigraph) on at most \( \sqrt{n} \) vertices.

We finally assume that each recursive call outputs an \( r \)-approximation of WMIS. Let \( j \in [d+1] \) be such that \( w(C_j \cap I) \geq \frac{1}{r+1} w(I) \) for \( I \) a heaviest independent set of \( G \) vertex-weighted by \( w \). Let \( J \subseteq \lceil \sqrt{n} \rceil \) be the indices of the \( P_i \)s that are intersected by \( C_j \cap I \), that is, \( J = \{ i : P_i \cap (C_j \cap I) \neq \emptyset \} \). For every \( i \in J \), set \( w_i = w(P_i \cap I) \). Each recursive call on some \( P_i \) with \( i \in J \), yields an independent set of weight at least \( \frac{w_i}{r+1} \), by assumption. Thus the weights that our algorithm puts on \( (G/P)[C_j] \) are such that it has an independent set of weight at least \( \frac{w(C_j \cap I)}{r+1} \). As we run an \( r \)-approximation on this graph, we get an independent set of weight at least \( \frac{w(C_j \cap I)}{r+1} \). Thus WMIS satisfies the assumptions of Lemma 22, and we conclude. \[\blacktriangleleft\]
4 Finding the suitable generalization: the case of Coloring

In this section, we deal with the Coloring problem. Unlike for WMIS, we cannot solely
resort to recursively calling our Coloring algorithm on smaller graphs. The right problem
generalization needs to be found for the inductive calls to work through, and it happens to
be \textsc{Set Coloring}.

In the \textsc{Set Coloring} problem, the input is a couple \((G, b)\) where \(G\) is a graph, and
\(b\) is a function assigning a positive integer to each vertex of \(G\). The goal is to find, for
each \(v \in V(G)\), a set \(S_v\) of at least \(b(v)\) colors such that \(S_u \cap S_v = 0\) whenever \(uv \in E(G)\),
and minimizing \(|\bigcup_{v \in V(G)} S_v|\). Let \(\chi_b(G)\) be the optimal value of \textsc{Set Coloring} for \((G, b)\).

Observe that \textsc{Coloring} corresponds to the case where \(b(v) = 1\) for every \(v \in V(G)\).

\textbf{Theorem 29.} \textsc{Set Coloring} (and hence \textsc{Coloring}) on \(n\)-vertex graphs \(G\) given with a
d\textsuperscript{\textcircled{2}}-sequence satisfies the assumptions of Lemma 24. In particular, this problem admits
\(a (d + 1)^{2d - 1}\)-approximation in time \(2^{O(d\cdot n^{\varepsilon} \log n)}\), for every integer \(q \geq 0\), and
\(a n^\varepsilon\)-approximation in polynomial-time for any \(\varepsilon > 0\), with \(d := c_2d + 2 \cdot 2^{d^2d + 2 + 4}\).

\textbf{Proof.} It is known [32] that \textsc{Set Coloring} can be solved using the inclusion-exclusion
principle in time \(O^\pi (\max_{v \in V(G)} b(v)^n) = 2^{O(n \log n)}\). We now prove that it satisfies the second
item of Lemma 22. We denote by \(A\) the \(r\)-approximation algorithm of the statement, which
we will use on instances of \textsc{Set Coloring}. In particular, we will call it at most \(\sqrt{n} + 1\) times, and will obtain at the end a \((d + 1) r^2\)-approximation on our input \((G, b)\) in polynomial
time.

We first apply Lemma 18 to get, in polynomial-time, a partition \(P = \{P_1, \ldots, P_{\lfloor \sqrt{n} \rfloor}\}\)
of \(V(G)\) whose parts have size at most \(d \sqrt{n}\) and such that \(R(G/P)\) has maximum degree
at most \(d\). For every \(i \in [\lfloor \sqrt{n} \rfloor]\), we use \(A\) to compute an \(r\)-approximated solution \(c_{P_i}\) of
\((G|P_i), b|_{P_i}\). We denote by \(b'\) the function which assigns, to each \(P_i\), the number of colors
of \(c_{P_i}\). We now compute, in polynomial-time, a proper \((d + 1)\)-coloring of \(R(G/P)\), which
defines the sets \(C_1, \ldots, C_{d+1}\). For each \(j \in [d + 1]\), we construct another \textsc{Set Coloring}
instance consisting of the graph \(H_j = (G/P)|C_j\) (recall that this trigraph has no red edge,
and can thus be seen as a graph), together with the function \(b'_{C_j}\). Again we use \(A\) to compute
an \(r\)-approximated solution on \((H_j, b'_{C_j})\). We denote by \(c_{H_j}\) this solution. Let \(G_j\) be the
subgraph of \(G\) induced by \(\bigcup_{P_i \in C_j} P_i\), and \(b_j\) the restriction of \(b\) to \(V(G_j)\). We now show how
to construct a solution \(c_j\) of \textsc{Set Coloring} to \((G_j, b_j)\) from \(c_{H_j}\) and all \(c_{P_i}\). Recall that
for every \(P_i \in C_j\), every \(v \in P_i\), we have that \(c_{P_i}(v)\) is a subset of \(\{1, \ldots, b'_{P_i}(v)\}\) of size at
least \(b(v)\), and that \(c_{H_j}(P_i)\) is a subset of size at least \(b'(P_i)\). Hence, for each \(P_i \in C_j\), one
choose an arbitrary bijection \(\tau\) from \(\{1, \ldots, b'(P_i)\}\) to \(c_{H_j}(P_i)\), and define to each vertex
\(v \in P_i\) the set \(c_j(v)\) as \(\{\tau(x) : x \in c_{P_i}(v)\}\).

By construction, this solution is a feasible one for the instance \((G_j, b_j)\). Let us prove
that it is an \(r^2\)-approximation of \(\chi_b(G_j)\). First, by definition of \(c_{H_j}\), our solution uses at
most \(r \cdot \chi_{b_{C_j}}(H_j)\) colors. Then, by definition of \(c_{P_i}\), for every \(P_i \in C_j\), we have \(b'_{C_j}(P_i) \leq
r \cdot \chi_{b_{P_i}}(G|P_i)\). Now, denote by \(\Gamma\) the function which assigns to each \(P_i \in C_j\) the number
\(\chi_{b_{P_i}}(G|P_i)\). We now use the following claim, whose proof is left to the reader.

\textbf{Claim 30.} Let \((G, b)\) be an instance of \textsc{Set Coloring}, and \(r \in \mathbb{R}_+\). It holds that
\(\chi_{r \cdot b}(G) \leq r \cdot \chi_b(G)\), where \(r \cdot b\) is the function which assigns \(r \cdot b(v)\) to each \(v \in V(G)\).

This implies \(\chi_{r \cdot b_{C_j}}(H_j) \leq r \cdot \chi_{r \cdot b}(H_j)\), and thus our solution uses at most \(r^2 \cdot \chi_{r \cdot b}(H_j)\) colors.

We now prove the following claim.
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Claim 31. \( \chi_r(H_j) \leq \chi_b(G_j) \).

Proof of the claim. Let \( c \) be an optimal solution for \((G_j, b_j)\). For every distinct \( P_i, P_r \in C_j \) such that \( P_i P_r \) is an edge of \( H_j \), it holds that there are all possible edges between \( P_i \) and \( P_r \) in \( G_j \) (by definition of the coloring \( C_1, \ldots, C_{d+1} \)), hence it holds that \( \cup_{v \in P_i} c(v) \) and \( \cup_{v \in P_r} c(v) \) have empty intersection. Moreover, by definition of \( \Gamma \), we have that \( \cup_{v \in P_i} c(v) \) is of size at least \( \Gamma(P_i) \), hence the function which assigns \( \cup_{v \in P_i} c(v) \) to each \( P_i \) is a feasible solution for \((H_j, \Gamma)\) using at most \( \chi_b(G_j) \) colors.

We now have in hand an \( r^2 \)-approximated solution of \((G_j, b_j)\) for every \( j \in [d+1] \), which can be turned into a \((d+1)r^2\)-approximated solution of \((G, b)\), as desired.

5 Edge-based problems: the case of Max Induced Matching

So far, we only considered problems where approximated solutions in each part \( P_i \) of a partition \( \mathcal{P} \) of \( V(G) \) of small width, and in some selected induced subgraphs of \((V(G/\mathcal{P}), E(G/\mathcal{P}))\), were enough to build an approximated solution for \( G \). We now handle problems for which a number of edges is to be optimized. Now all competitive solutions can integrally lie in between pairs of parts \( P_i, P_j \) linked by a black or a red edge in \( G/\mathcal{P} \). This complicates matters, and forces us to be competitive there as well, naturally splitting the algorithms into three subroutines.

We present the algorithms for MAX SUBSET INDUCED MATCHING where one is given, in addition to the input graph \( G \) (possibly with edge weights), a subset \( Y \subseteq E(G) \), and the goal is to find a heaviest induced matching \( S \) of \( G \) such that \( S \subseteq Y \). Then MAX INDUCED MATCHING is the particular case when \( Y = E(G) \). Of course, we could solely use the edge weights to emulate \( Y \) (by giving negative weights to all the edges in \( E(G) \setminus Y \)). We believe this formalism is slightly more convenient for the reader to quickly and explicitly identify where our algorithm is seeking mutually induced edges.

Since the case of MAX INDUCED MATCHING is more involved than were the treatment of MIS and COLORING, we again split the arguments into the design of a subexponential-time constant-approximation algorithm (Lemma 34) followed by how this algorithm meets the requirements of Lemma 22 (Theorem 33).

Lemma 32. Assume every input \((G, Y)\) is given with a \( d' \)-sequence of the \( n \)-vertex, edge-weighted by \( w \), graph \( G \). We set \( d := c_{d'} \cdot 2^{4w} + 4 \), and \( s := 2^{4w} + 4 \). MAX SUBSET INDUCED MATCHING can be \( O(d^2) \)-approximated in time \( 2^{O(s\sqrt{n})} \) on these inputs.

Proof. Again, by Lemma 17, we start by computing in polynomial time a partition of \( V(G) \), \( \mathcal{P} = \{P_1, \ldots, P_{\sqrt{n}}\} \), of parts with size at most \( s\sqrt{n} \) and such that \( \mathcal{R}(G/\mathcal{P}) \) has maximum degree at most \( d \).

We \((d+1)\)-color \( \mathcal{R}(G/\mathcal{P}) \), which defines a coarsening \( \{C_1, \ldots, C_{d+1}\} \) of \( \mathcal{P} \). We also distance-2-edge-color \( \mathcal{R}(G/\mathcal{P}) \) with \( z = 2(d - 1)d + 1 \) colors, that is, properly (vertex-)color the square of its line graph. Observe that \( z - 1 \) upperbounds the maximum degree of the square of the line graph of \( \mathcal{R}(G/\mathcal{P}) \). This partitions the edges of \( \mathcal{R}(G/\mathcal{P}) \) into \( \{E_1, \ldots, E_z\} \). For each red edge \( e = P_i P_j \in \mathcal{R}(G/\mathcal{P}) \), we denote by \( p(e) \) the set \( P_i \cup P_j \). We also set \( X_h = p(E_h) = \bigcup_{e \in E_h} p(e) \) for each \( h \in [z] \).

Let \( M \subseteq Y \) be a fixed (unknown) heaviest induced matching of \( G \) contained in \( Y \). Let \( M_c, M_r, M_b \) partition \( M \), where \( M_c \) (as vertex) consists of the edges of \( M \) with both

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5 The improvement based on clustered coloring slightly departed from that simple scheme.
endpoints in a same $P_i$, $M_r$ (as red) corresponds to edges of $M$ between some $P_i$ and $P_j$ with $P, P_j \in R(G/P)$, and $M_b$ (as black), the edges of $M$ between some $P_i$ and $P_j$ with $P, P_j \in E(G/P)$. We compute three induced matchings $N_v, N_r, N_b \subseteq Y$ of $G$, capturing a positive fraction of $M_v, M_r, M_b$, respectively. Figure 3 gives the intuition of the procedures which determine each of these approximated solutions.

**Computing $N_v$.** For every integer $1 \leq i \leq \lceil \sqrt{n} \rceil$, we compute a heaviest induced matching in $G[P_i]$ contained in $Y$, say $S_i$, in time $2^{O_d(\sqrt{n})}$. For each $j \in [d + 1]$, let $H_j$ be the graph $(G/P)[C_j]$ with every vertex $P_i \in C_j$ weighted by $w(S_i)$. We compute a heaviest independent set $I_j$ in $H_j$, also in time $2^{O_d(\sqrt{n})}$.

Let $R_i$ be the induced matching $\{e \in S_i : P_i \in I_j\}$. It is indeed an induced matching in $G$ contained in $Y$, since each $S_i$ is so, there is no red edge in $(G/P)[C_j]$, and $I_j$ is an independent set of $H_j$. The solution $N_v$ is then a heaviest among the $R_i s$.

**Computing $N_r$.** For each $e = P_i P_j \in R(G/P)$, we compute a heaviest induced matching $S'_r$ in $G[p(e)] = G[P_i \cup P_j]$ among those that are included in $Y$ and have only edges with one endpoint in $P_i$ and the other endpoint in $P_j$. This takes times at most $\frac{2^{O_d(\sqrt{n})}}{2^{O_d(\sqrt{n})}} = 2^{O_d(\sqrt{n})}$ by trying out all vertex subsets, since $|P_i \cup P_j| \leq 2s \sqrt{n}$. For each $h \in [z]$, let $H'_h$ be the graph $(G/P)[\{P_i : P_i \text{ is incident to an edge } e \in E_h\}]$ and the red edges $e \in E_h$ are turned black and get weight $w(S'_r)$. We compute a heaviest induced matching $I'_h$ in $H'_h$ among those included in $E_h$, in time $2^{O_d(\sqrt{n})}$. Note here that we changed the prescribed set of edges $Y$ to $E_h$.

Let $R'_h$ be the induced matching $\{f \in S'_r : e \in I'_h\} \subseteq Y$ of $G$. Indeed, each $S'_r \subseteq Y$ is an induced matching, and there is no red edge between an endpoint of $e \in I'_h$ and an endpoint of $e' \neq e \in I'_h$ (since $E_h$ is a color class in a distance-2-edge-coloring of $R(G/P)$), nor a black edge (by virtue of $I'_h$ being an induced matching of $H'_h$). The solution $N_r$ is then a heaviest among the $R'_h s$.

Figure 3 Illustration of how to determine the induced matching $N_v$, $N_r$, and $N_b$ (in that order, from left to right).
Computing $N_b$. Observe first that an induced matching of $G$ can only contain at most one edge between $P_1$ and $P_2$ when $P_1, P_2 \in E(G/P)$. Thus in the graph $(V(G/P), E(G/P))$, we give weight $\max \{w(f) : f = uv \in Y, u \in P_1, v \in P_2\}$, with the convention that $\max \emptyset = -1$, to each edge $e = P_1, P_2 \in E(G/P)$, call $G'$ the resulting edge-weighted graph, and denote by $m_G(e)$ an edge $f \in Y$ realizing this maximum. We compute a heaviest induced matching $S$ of $G'$ included in $E(G')$, in time $2^{O(d(\sqrt{n}))}$. Let $H_S$ be the graph with vertex set $S$, and an edge between $e$ and $e'$ whenever there is a red edge in $G/P$ between an endpoint of $e$ and an endpoint of $e'$. As $H_S$ has degree at most $2d$, it can be $2d + 1$-colored; let $T_1, \ldots, T_{2d+1}$ the corresponding color classes.

For each $i \in [2d + 1]$, let $R_i'$ be the induced matching $\{m_G(e) : e \in T_i\} \subseteq Y$ of $G$. Indeed, $S$ is an induced matching in the black graph of $G/P$, and the underlying vertices of $T_i$ do not induce any red edge in $G/P$, by design. The solution $N_r$ is then a heaviest among the $R_i'$'s.

We finally output a heaviest set among $N_v, N_r, N_b$. The overall running time is $2^{O(d(\sqrt{n}))}$ as we make a polynomial number of calls to (exhaustive) subroutines on graphs with $O(d(\sqrt{n}))$ vertices, and color in linear time $O(n^2)$-vertex graphs of maximum degree $\Delta$ with $\Delta + 1$ colors. We already argued that $N_v, N_r, N_b \subseteq Y$ are all induced matchings in $G$, thus so is our output.

We shall just show that we meet the claimed approximation factor. First, one can observe $w(N_v) \geq w(M_r)/d$. Second, at least a $\frac{1}{2}$ fraction of the weight of $M_r$ intersects some fixed $E_i$ (with $i \in [z]$). Let $\mathcal{J}$ be the parts of $\mathcal{P}$ intersected by $M_r \cap X_i$. As there cannot be a black edge between two parts of $\mathcal{J}$ (otherwise $M_r$ is not an induced matching as defined), our algorithm indeed computes an induced matching of $G[X_i]$ included in $Y$ of weight at least $w(M_r \cap X_i)$. Hence $w(N_v) \geq w(M_r)/2$.

Third, we already argued that an induced matching in $G'$ corresponds to an induced matching in the black graph of $G/P$. Thus at least one of the $R_i'$ (with $i \in [2d + 1]$) contains at least a $\frac{1}{2d + 1}$ fraction of the weight of $M_b$. Therefore $w(N_b) \geq w(M) \geq 2^{1/(2d+1)}$.

Finally the output induced matching has at least weight

$$\frac{w(M)}{3 \cdot \max(d + 1, \beta, 2d + 1)} = \frac{w(M)}{3 \cdot 2d + 1} = \frac{w(M)}{3(2d - 1)d + 1}. \quad \blashed{\Box}$$

**Theorem 33.** **Max Subset Induced Matching** on an $n$-vertex graph $G$, edge-weighted by $w$, with prescribed set $Y \subseteq E(G)$, and given with a $d'$-sequence, satisfies the assumptions of Lemma 22. In particular, this problem admits

- a $(d + 1)^{2^{d-1}}$-approximation in time $2^{O(d(\sqrt{n}))}$, for every integer $q \geq 0$,
- an $n^{\frac{2}{d'}}$-approximation in polynomial-time $O_{d,c}(1)\log n^{O(1)} n \cdot n^{O(1)}$, for any $\varepsilon > 0$, and
- a log $n$-approximation in time $2^{O_d(n^{\frac{2}{d'}+\frac{\log n}{t(\frac{1}{\varepsilon})}})}$,

with $d := 2d + 2 \cdot 2^{2d+1} \cdot 2^{d+1} \cdot 2^{d+1}$. 

**Proof.** The exhaustive algorithm (trying out all vertex subsets and checking whether they induce a matching included in $Y$) solves MAX SUBSET INDUCED MATCHING in time $\sigma^{O(n)}$.

Thus we show MAX SUBSET INDUCED MATCHING satisfies the second item of Lemma 22, as witnessed by Lemma 34 where subcalls are dealt with recursively. We set $c_2 \geq 1$ as the required exponent to turn a $d'$-sequence into a neatly divided matrix of $M_{n, 2d'+2}$, and compute the various needed colorings, the appropriate $\frac{1}{2} < c_1 < 1$, and $2 < c_3 < 3$, and $f(d) = 2d \geq 1$ with $s := 2^{d+1}$.

In computing $N_v$, the algorithm makes $\lceil \sqrt{n} \rceil$ recursive calls and $d + 1$ calls to WEIGHTED MAX INDEPENDENT SET on induced subgraphs of $G$. All of these induced subgraphs are on less than $f(d)\sqrt{n}$ vertices. Computing $N_r$ makes at most $\frac{\sqrt{n}d}{2}$ recursive calls on induced
subgraphs of $G$ with at most $f(d)\sqrt{n}$ vertices, followed by at most $2(d - 1)d + 1$ recursive calls on full cleanups of induced subtrigraphs of $G/P$ with at most $\sqrt{n}$ vertices (in fact, one can observe that the latter recursive calls happen to also be on induced subgraphs of $G$).

Finally, computing $N_b$ makes one recursive call to a full cleanup of $G/P$ on $[\sqrt{n}]$ vertices.

In summary, we make $O_d(\sqrt{n})$ recursive calls or calls to another problem WMIS (which already satisfies Lemma 22 with better constants) on induced subgraphs of $G$ or full cleanups of (the whole) $G/P$, each on $O_d(\sqrt{n})$ vertices. Hence, by Lemma 18, the induction applies.

We check that getting $r$-approximations on every subcall allows to output a global $3(2(d - 1)d + 1)r^2$-approximation. For that we argue that $N_r$ (resp., $N_r$, $N_b$) is a $(2(d - 1)d + 1)r^2$-approximation of $M_r$ (resp., $M_r$, $M_b$). The fact that $N_r$ is a $(d + 1)r^2$-approximation (hence a $(2(d - 1)d + 1)r^2$-approximation, since we assume that $d \geq 1$) of $M_r$ directly follows Theorem 28.

We now show that $N_r$ is a $(2(d - 1)d + 1)r^2$-approximation of $M_r$. Let $h \in [z] = [2(d - 1)d + 1]$ be an index maximizing $w(M_r \cap E(G[X_h]))$. Thus $w(M_r \cap E(G[X_h])) \geq w(M_r)/4d(d - 1)d - 1$.

Let $F_h \subseteq E_h$ be the edges $e = P_iP_j$ of $\mathcal{R}(G/P)$ that are inhabited by $M_r$ (i.e., $M_r$ contains at least one edge between $P_i$ and $P_j$). Note that our algorithm makes an $r$-approximation of the optimum such solutions on $p(e)$ (selecting only edges between $P_i$ and $P_j$). Thus the $r$-approximation on $H_h'$ yields the desired $(2(d - 1)d + 1)r^2$-approximation $N_r$.

Finally, one can easily see that $N_b$ is a $(2d + 1)r$-approximation of $M_b$ (note, here, the absence of a 2 in the exponent of $r$).

\section{Technical generalizations}

\subsection{Mutually Induced $H$-packing}

In this section we present a far-reaching generalization of the approximation algorithms for \textsc{Max Independent Set} and \textsc{Max Induced Matching}. For any fixed graph $H$, let \textsc{Mutually Induced $H$-packing} be the problem where one seeks a largest collection of mutually induced copies of $H$ in the input graph $G$, that is, a largest set $S$ such that $G[S]$ is a disjoint union of (copies of) graphs $H$. We get similar approximation guarantees for \textsc{Mutually Induced $H$-packing}, for any connected graph $H$. Observe that \textsc{Max Independent Set} and \textsc{Max Induced Matching} are the special cases when $H$ is a single vertex and a single edge, respectively.

We in fact approximate a technical generalization that we call \textsc{Annotated Mutually Induced $H$-packing}. The input is a tuple $(G, w, z, \gamma, \gamma')$ where $G$ is a graph, $w : V(G)^{|V(H)|} \rightarrow \mathbb{Q}$ is a weight function over the tuples \textit{without repetition} of $V(G)$ of size $|V(H)|$ (that we will use to keep track of the number of mutually induced copies \textit{within} a given tuple of vertices of $G$), $z$ is an integer between 1 and $|V(H)|$, $\gamma : V(G) \rightarrow [z]$ is a labeled partition of $V(G)$ into $z$ classes, and $\gamma' : V(H) \rightarrow [z]$ is a labeled partition of $V(H)$ into $z$ classes. Note that the \textsc{Mutually Induced $H$-packing} is obtained when $w(Z) = |G[Z]|$ is isomorphic to $H$ (where $[\cdot]$ is the Iverson bracket, i.e., taking value 1 if the property it surrounds is true, and 0 otherwise) and $z = 1$ (which forces the value of $\gamma$ and $\gamma'$). The goal is to find a subset $S$ such that

- $G[S]$ is a disjoint union of copies of $H$,
- there is an isomorphism between each copy $C$ of $H$ (in $S$) and $H$ which preserves $\gamma, \gamma'$, i.e., every vertex $v$ of $C$ is mapped to a vertex $v' \in V(H)$ with $\gamma(v) = \gamma'(v')$, and
- $\sum_{C \text{ copy of } H \text{ in } S} w(V(C))$ is maximized.

We will need the notion of \textit{compatible trigraphs} of a (labeled) graph. Given a graph $H$,
we call compatible trigraph of $H$ any trigraph on at most $|V(H)|$ vertices obtained by turning some (possibly none) black edges or non-edges of trigraph $H/Q$ (for any fixed choice of a partition $Q$ of $V(H)$) into red edges. In other words, a compatible trigraph $H'$ of $H$ is such that there is a cleanup $H''$ of $H'$ that is also a quotient trigraph of $H$. Note that the number of compatible trigraphs of an $h$-vertex graph $H$ is upperbounded by $B_h \cdot 2\binom{h}{2} = 2^{O(h^2)}$, where $B_h$ is the $h$-th Bell number, which counts the number of partitions of a set of size $h$.

Given a graph $G$ vertex-partitioned by $\mathcal{P}$ and a trigraph $H$, a subset $S \subseteq V(G)$ is said cut by $\mathcal{P}$ along $H$ if $G[S]/\mathcal{P}$ is isomorphic to $H$. By extension, the copy of $G[S]$ in $G$ (induced by $S$) is also said cut by $\mathcal{P}$ along $H$.

\textbf{Lemma 34.} For any connected graph $H$, \textsc{Annotated Mutually Induced $H$-packing}, when every input $(G, w, z, \gamma, \gamma')$ is given with a $d'$-sequence of the $n$-vertex graph $G$, satisfies the assumptions of Lemma 22. In particular, this problem admits

\begin{itemize}
  \item a $d'^{O_h(2^q)}$-approximation in time $2^{O_{d,h,q}(n^2-q)}$, for every integer $q \geq 0$,
  \item an $n^{c}$-approximation in polynomial-time $O_{d}(1) \cdot n^{O_{d,h}(1)}$, for any $\varepsilon > 0$,
\end{itemize}

with $h = |V(H)|$, and $d := c_{d'} + 2 \cdot 2^{4c_{d'} + 2 + 4}$.

\textbf{Proof.} As the first item of Lemma 22 is satisfied, we describe an algorithm that fulfills the requirement of its second item. We proceed by induction on the number of vertices of $H$.

Thus we can assume that \textsc{Annotated Mutually Induced $J$-packing}, with $J$ a connected graph on less vertices than $H$, satisfies Lemma 22. We already did the base case of the induction, which was \textsc{Weighted Max Independent Set}.

\textbf{Algorithm.} Again, by Lemma 18, we start by computing in polynomial time a partition of $V(G)$, $\mathcal{P} = \{P_1, \ldots, P_{\lceil \sqrt{n} \rceil}\}$, of parts with size at most $d\sqrt{n}$ and such that $\mathcal{R}(G/\mathcal{P})$ has maximum degree at most $d$. Let $S$ be a fixed (unknown) heaviest (with respect to $w$) mutually induced $H$-packing of $G$ preserving $\gamma, \gamma'$.

For every compatible trigraph $H'$ of $H$, we look for mutually induced copies of $H$ in $G$ cut by $\mathcal{P}$ along $H'$, and preserving $\gamma, \gamma'$. As the number of compatible trigraphs of $H$ is $2^{O(h^2)}$, a $1/2^{O(h^2)}$ fraction of the weight of $S$ is made of mutually induced copies of $H$ which are cut by $\mathcal{P}$ along a fixed compatible trigraph $H'$. We now focus on this particular “run.”

We distinguish two cases:

\begin{itemize}
  \item $(A)$ $H'$ has at least one black edge, or
  \item $(B)$ $H'$ has no black edge.
\end{itemize}

As $H$ is connected, the total graph of $H'$ is also connected. Indeed, switching some edges or non-edge to red edges in the quotient trigraph of $H$ cannot disconnect the total graph, which can only gain edges. Thus in case (A), every red component of $H'$ has at least one incident black edge, and in case (B), $H'$ has a single red component (and no black edge).

In general, we want to individually pack red components of $H'$ (first type of recursive calls in smaller induced subgraphs of $G$), then combine those red components by connecting them with the right pattern of black edges (second type of recursive calls in the total graph of $G/\mathcal{P}$). Handling both cases (A) and (B) in an unified way runs into the technical issue that the weight function may destroy our combined solutions in an uncontrollable manner.

The case distinction works as a win-win argument. In case (A), due to the presence of a black edge in $H'$, we can pack at most one mutually induced copy of $H$ within any fixed subtrigraph of $G/\mathcal{P}$ matching $H'$. We thus exempt ourselves from the first type of recursive calls. In case (B), we do need the two types of recursive calls (as in WMIS), but the first type is done on the whole $H$. Thus the current weight function (on $h$-tuples) is informative enough.
**Case (A).** The essential element here is to build a new weight function \( w' \) on the \( h' \)-tuples of the total graph \( T(G/P) \), with \( h' := |V(H')| \). For every injective map \( \iota : V(H') \to P \) inducing a trigraph isomorphism and preserving \( \gamma, \gamma' \), for every ordering of \( \iota(V(H')) \) into an \( h' \)-tuple \((P_1, \ldots, P_{h'})\), we set
\[
w'(P_1, \ldots, P_{h'}) := \max\{w(v_1^1, v_1^2, \ldots, v_1^{a_1}, \ldots, v_h^1, v_h^2, \ldots, v_h^{a_h}) : v_1^1, v_1^2, \ldots, v_h^{a_h} \in P_1, \ldots, v_1^1, v_2^1, \ldots, v_h^{a_h} \in P_{h'}\},
\]
and \( G'[v_1^1, v_1^2, \ldots, v_1^{a_1}, \ldots, v_h^1, v_h^2, \ldots, v_h^{a_h}] \) is isomorphic to \( H \).

Indeed as we previously observed, in case (A), at most one mutually induced copy of \( H \) respecting the cut along \( H' \) can be packed in the subgraph of \( G \) induced by the vertices of \( \iota(V(H')) \). (In the definition of \( w' \), we can further impose that \( a_i \) matches the number of vertices of \( H \) in the corresponding part of \( H' \) but this is not necessary.)

All the \( h' \)-tuples not getting an image by \( w' \) in the previous loop (realized in time \( n^{O(h)} \)) are assigned the value 0. We then make a recursive call to **Annotated Mutually Induced \( T(H') \)-packing** on input \((T(G/P), w', 1, \gamma_0, \gamma'_0)\) where we recall that \( T(\iota) \) is the total graph, and \( \gamma_0, \gamma'_0 \) are the constant 1 functions.

**Case (B).** For every injective map \( \iota : V(H') \to P \) inducing a trigraph isomorphism and preserving \( \gamma, \gamma' \), we make a recursive call to **Annotated Mutually Induced \( H \)-packing** with input \((G_1 = G[\bigcup_{P \in \iota(V(H'))} P], w, h, \gamma, \gamma'_1)\) where two vertices get the same label by \( \gamma \) if and only if they have the same label by \( \gamma \) and lie in the same \( P \in \iota(V(H')) \), and \( \gamma'_1 \) forces to a vertex \( v' \in X \in V(H') \) of \( H \) the same label given to the vertices \( v \in \iota(X) \) such that \( \gamma'_1(v') = \gamma(v) \). Informally \( \gamma, \gamma'_1 \) forces the recursive call to commit to the map \( \iota \) and the former functions \( \gamma, \gamma' \).

Each such recursive call yields a mutually induced packing of \( H \). Since the red graph of \( G/P \) has degree at most \( d \), we can color the (ordered) tuples of \( P \) of length up to \( h \) and inducing a connected subgraph of \( R(G/P) \) with at most \( p(h, d) = h^{2d^h} \cdot d^{2h} \cdot h! + 1 \) colors such that every color class consists of disjoint tuples pairwise not linked by a red edge in \( G/P \). Indeed the claimed number of colors minus 1 upperbounds, in \( R(G/P) \), the number of connected tuples of length up to \( h \) that can touch (i.e., intersect or be adjacent to) a fixed connected tuple of length up to \( h \). One color class contains a fraction \( 1/p(h, d) \) of the weight of the optimal solution \( S \) (subject to the same constraints). Running through all color classes \( j \) (and focusing on one containing a largest fraction of the optimum), we define a weight function \( w' \) on the \( h' \)-tuples of \( T(G/P) \), with \( h' = |V(H')| \), by giving to a tuple the weight returned by the corresponding recursive call whenever it is part of color class \( j \), and weight 0 otherwise. We then make a recursive call to **Annotated Mutually Induced \( T(H') \)-packing** on input \((T(G/P), w', 1, \gamma_0, \gamma'_0)\) where we recall that \( T(\iota) \) is the total graph, and \( \gamma_0, \gamma'_0 \) are the constant 1 functions.

We output a heaviest solution among all runs. We now check that the algorithm is as prescribed by Lemma 22.

**Number of recursive calls.** We make at most \( 2^{O(h^2)} \cdot h \cdot |V(G/P)|^h = n^{O(h)}(1) \) recursive calls to **Annotated Mutually Induced \( H \)-packing**, and at most \( p(h, d) + 1 = O_{d,h}(1) \) recursive calls to **Annotated Mutually Induced \( T(H') \)-packing**. Hence there is a constant \( c_1 \) (function of \( d \) and \( h \)) such that the number of calls is bounded by \( n^{c_1} \).

**Nature and size of the inputs of the recursive calls.** Both \( H \) and \( T(H') \) have strictly less vertices than \( H \) or are equal to \( H \). Thus the induction on \( h \) applies. Besides, \( G[\bigcup_{P \in \iota(V(H'))} P] \) is an induced subgraph of \( G \) of size at most \( h \cdot d \sqrt{n} = O_{d,h}(1) \cdot \sqrt{n} \), and \( T(G/P) \) is a full cleanup of \( G/P \) of size at most \( \lfloor \sqrt{n} \rfloor \).
Running time. Outside of the recursive calls, one can observe that our algorithm takes times \( O_{d,h}(1) \cdot n^{O_r(1)} \). Hence there is a constant \( c_2 \) (function of \( d \) and \( h \)) such that the running time of that part is bounded by \( n^{c_2} \).

Correctness and approximation guarantee. As all the recursive calls are on induced subgraphs of \( G \) or of the total graph \( T(G/P) \), we return a mutually induced collection of graphs of the size of \( H \). All these graphs are indeed induced copies of \( H \) since the weight function prevents the false positives of copies of \( H \) in the total graph \( T(G/P) \) but not in \( G \) (these tuples are given weight 0). Finally it can be checked that the returned solution has weight a fraction \( (2^{O(h^2)} \cdot \max(r,p(h,d)r^2))^{-1} \) of the optimum, which can also be seen as a \( d^{-2} \cdot r^2 \)-approximation for some constant \( c_3 \) depending on \( d \) and \( h \).

6.2 Independent induced packing of stars and forests

The techniques employed to design approximations algorithms for \textsc{Max Subset Induced Matching} can be extended in order to tackle more general problems. In particular, we show in this section a generalization of Theorem 33 for \textsc{Max Edge Induced Star Forest} and \textsc{Max Edge Induced Forest}. These two problems stand as the version of \textsc{Mutually Induced} \( H \)-\textsc{packing} where \( H \) is respectively either the infinite family of stars or trees.

On the one hand, \textsc{Max Edge Induced Star Forest} asks, given a graph \( G \) and a subset \( Y \subseteq E(G) \), for a collection of induced stars on \( G \), made up of edges of \( Y \) only, maximizing the number of edges (or leaves).

\begin{tabular}{|l|}
\hline
\textbf{\textsc{Max Edge Induced Star Forest}} \\
\textbf{Input:} Graph \( G \), subset \( Y \subseteq E(G) \) \\
\textbf{Output:} Collection \( (A_i)_{i \in [k]} \) of induced stars on \( G \), made up of edges in \( Y \) only, such that there is no edge between \( A_i \) and \( A_j \), for any \( i \neq j \in [k] \), which maximizes the number of edges. \\
\hline
\end{tabular}

On the other hand, given the same input, \textsc{Max Edge Induced Forest} asks for an induced forest \( F \) on \( G \) with the largest set of edges.

We would like to emphasize the fact that the objective function of both problems counts the number of \textit{edges} in the solution, instead of \textit{vertices}, as it is often the case in the literature when looking for a collection of stars or trees in a graph. The reason for this is because an approximated solution for these vertex versions can be obtained from an approximated solution of \textsc{Weighted Max Independent Set} (since any independent set is a star forest, and any forest is a bipartite graph).

Observe moreover that a solution of \textsc{Max Edge Induced Forest} can be 3-approximated with a solution of \textsc{Max Edge Induced Star Forest}. Indeed, the edge set of any tree can be partitioned into three distance-2-edge colors, which consist of a collection of stars. Therefore, the induced forest \( F \) can be partitioned into three collections of induced stars. In the remainder, we design approximation algorithms for \textsc{Max Edge Induced Star Forest}, and directly deduce results for \textsc{Max Edge Induced Forest}.

In the remainder, we propose approximation algorithms for \textsc{Max Edge Induced Star Forest}. We provide in particular a \( n^2 \)-approximation algorithm for \textsc{Max Edge Induced Star Forest}, running in polynomial time.

We need to find the suitable generalization of \textsc{Max Edge Induced Star Forest}, as it was done for \textsc{Coloring} in Section 4. We call this problem \textsc{Max Leaves Induced Star Forest}. Now, a weight function on vertices is added to the input, and we seek a collection of mutually induced stars with maximum weight, the weight of a star being the sum of the weights of its leaves (that is, the weight of the root is omitted).
**Max Leaves Induced Star Forest**

**Input:** Graph $G$, weights $w_V : V \to \mathbb{N}$, subset $Y \subseteq E(G)$

**Output:** Collection $(A_i)_{i \in [k]}$ of induced stars on $G$ with root $r_i$, $A_i = \{r_i, s_{i1}^1, \ldots, s_{ik_i}^i\}$, made up of edges in $Y$ only, with no edge between $A_i$ and $A_j$, for any $i \neq j \in [k]$, maximizing

$$\sum_{i=1}^k w_V(A_i) = \sum_{i=1}^k \sum_{j=1}^{L_i} w(s_{ij}^i)$$

We prove that **Max Leaves Induced Star Forest** follows the framework proposed in Lemma 22. We begin with the design of a subexponential-time algorithm approximating a solution of **Max Leaves Induced Star Forest** with a ratio function of twin-width.

**Lemma 35.** Assume every input of **Max Leaves Induced Star Forest** is given with a $d'$-sequence of the $n$-vertex $G$, and $d := e^{2\varepsilon^2 + 2 \cdot 2^{4\varepsilon^2 + \varepsilon^4}}$. **Max Leaves Induced Star Forest** can be $O(d^2)$-approximated in time $2^{O(d^2 \sqrt{n})}$ on these inputs.

**Proof.** We compute in polynomial time a partition of $V(G)$, $P = \{P_1, \ldots, P_{\sqrt{n}}\}$, of parts with size at most $d' \sqrt{n}$ and such that $R(G/P)$ has maximum degree at most $d$, by Lemma 17.

As in Lemma 19, we $(d+1)$-color $R(G/P)$, which defines a coarsening $\{C_1, \ldots, C_{d+1}\}$ of $P$. Moreover, we distance-2-edge-color $R(G/P)$ with $z = 2(d-1)d + 1$ colors. This partitions the edges of $R(G/P)$ into $\{E_1, \ldots, E_z\}$. For each red edge $e = P_iP_j \in R(G/P)$, we denote by $p(e)$ the set $P_i \cup P_j$.

Let $A = \bigcup_{i=1}^k A_i$ be the union of all stars present in an optimum solution of **Max Leaves Induced Star Forest** in $G$. We have $A \subseteq Y$. Let $A_v, A_r, A_b$ partition $A$, where $A_v$ contains the edges of $A$ with both endpoints in the same $P_i$, $A_r$ corresponds to edges of $A$ between some $P_i$ and $P_j$ with $P_iP_j \in R(G/P)$, and $A_b$ the edges of $A$ between some $P_i$ and $P_j$ with $P_iP_j \in E(G/P)$. The set of edges $A_v$ (resp. $A_r$, $A_b$) still form a collection of mutually induced stars. At least one over the three solutions produced by the partition $A_v, A_r, A_b$ gives us a 3-approximation for this problem. Our algorithm consists of computing three solutions for **Max Leaves Induced Star Forest** of $G$, capturing a positive fraction of $A_v, A_r, A_b$, respectively.

**Computing a $d+1$-approx for $A_v$.** **Construction.** For every integer $1 \leq i \leq \lceil \sqrt{n} \rceil$, we compute an optimum solution for **Max Leaves Induced Star Forest** in $G[P_i]$ contained in $Y$, say $S_i$, in time $2^{O(d \sqrt{n})}$. This can be achieved with guesses of the vertices in $P_i$, as $|P_i| \leq d' \sqrt{n}$.

Then, we focus on each color $C_j$ of $R(G/P)$, for $j \in [d+1]$. There is no red edge in $H_j = (G/P)[C_j]$. We compute a heaviest independent set $I_j$ in $H_j$ where the parts $P_i$ are weighted by the edge weight of $S_i$. Let $R_j$ be the union of all optimum solutions for **Max Leaves Induced Star Forest** on all $P_i$ belonging to $I_j$. The solution returned is the maximum over all $R_j$.

**Approximation ratio.** Let $A'_i$ be the subset of $A_v$ made up of edges belonging to parts of $C_j$. There is no red edge between two parts of $C_j$, therefore their neighborhood consists of either full adjacency or full non-adjacency. As a consequence, a maximum-weighted collection of stars in $C_j$ with edges inside parts intersects parts which are pairwise non-adjacent in $(G/P)[C_j]$, otherwise the stars are not mutually induced. Consequently, this justifies that the set $R_j$ returned for each $C_j$ is a maximum-weighted collection of stars in $C_j$ made up of edges inside parts. In summary, the weight of each collection $R_j$ is greater than the weight of $A'_j$. As $j \in [d+1]$, a heaviest collection among all $R_j$s is a $d+1$-approximation of $A_v$.

**Computing an $O(d^2)$-approx for $A_r$.** **Construction.** For each $e = P_iP_j \in R(G/P)$, we
compute an optimal solution for MAX LEAVES INDUCED STAR FOREST in $G[p(e)] = G[P_i \cup P_j]$ among those that are included in $Y$ and have only edges with one endpoint in $P_i$ and the other endpoint in $P_j$. Said differently, we determine a maximum-weighted collection of induced stars in $G[p(e)]$ over $Y$ with a root on one side (for example, $P_i$) and all leaves on the other side ($P_j$). This costs at most $2^{O(d(\sqrt{n}))}$ by trying out all vertex subsets, since $|P_i \cup P_j| \leq 2d\sqrt{n}$. The set of vertices of the solution returned on $G[p(e)]$ is denoted by $B_e \subseteq p(e)$.

For each $h \in [z]$, let $H'_h$ be the trigraph $(G/P)\{P_i : P_i$ is incident to an edge $e \in E_h]\$. The red edges of $H'_h$ form an induced matching on the red graph of $H'_h$ as they are at distance 2 in $G/P$. We associate with any edge $e \in E_h$ the edge weight of $B_e$. Then, we turn the red edges of $H'_h$ in black: let $H''_h$ be the graph obtained. We solve MAX SUBSET INDUCED MATCHING on $H''_h$ by restricting it to edges of $E_h$ (which plays the role of $Y$): this is achieved in $2^{O(\sqrt{n})}$ as $|V(H''_h)| \leq \sqrt{n}$. Let $I''_h$ be a maximum-weighted induced matching obtained. For each $h \in [z]$, we obtain the union $R_h$ of all $B_e$, $e \in I''_h$: $R_h = \bigcup_{e \in I''_h} B_e$. We return an $R_h$ which maximizes the total edge weight, among all $h \in [z]$.

**Approximation ratio.** Let $A'^h$ be the subset of $A_h$ made up of edges being part of red edges $E_h$ in $G/P$, for $h \in [z]$. As the edges of $E_h$ form an induced matching in $R(G/P)$, the union of solutions of MAX LEAVES INDUCED STAR FOREST on graphs $G[p(e)]$ with $e \in E_h$ can only be connected through black edges of $G/P$. Furthermore, two collections of stars over $G[p(e)]$ and $G[p(f)]$ are necessarily not mutually induced if there is a black edge between an endpoint of $e$ and an endpoint of $f$. Consequently, $R_h$ gives a maximum-weighted collection of mutually induced stars over $E_h$ and its weight is at least the weight of $A'^h$. The maximum-weighted collection over all $R_h$ gives a $z$-approximation, as $h \in [z]$.

**Computing a $2d + 1$-approx for $A_h$.** Construction. For each part $P_i$, we solve WEIGHTED MAX INDEPENDENT SET on $G[P_i]$ with weight function $w_V$. Let $I(P_i)$ be the independent set returned and $w(P_i)$ its weight. We focus now on graph $G' = (V(G/P), E(G/P))$, made up of the black edges of $G/P$, and solve MAX LEAVES INDUCED STAR FOREST on it with weights $w(P_i)$. As $|V(G')| \leq \sqrt{n}$, this is achieved in $2^{O(\sqrt{n})}$.

Let $(B_h)_{h \in [k]}$ be the collection of stars returned, $B_h = \{R_h, S_h^1, \ldots, S_h^{L_h}\}$ and $B \in E(G')$ be the set of edges belonging to this collection. Based on the bounded maximum red degree of $G/P$, we determine a $O(d)$-partition of the edges of $B$, in order to produce collections of mutually induced stars. Let $H^*$ be the graph where each edge $e$ in the collection $(B_h)_{h \in [k]}$ is represented with a vertex and two of them $e, f$ are adjacent if and only if there is a red edge in $G/P$ connecting an endpoint of $e$ with an endpoint of $f$. This graph has degree at most $2d$, so it can be $2d + 1$-colored: let $T_1, \ldots, T_{2d+1}$ be the corresponding color classes. Any set of edges $T_j$ gives us a collection of mutually induced stars on trigraph $G/P$, in the sense that there is neither a black nor a red edge between two stars.

We fix some color class: say $T_1$ w.l.o.g. Let $(B^*_h)$ be the collection of stars produced by $T_1$, where $B^*_h = \{R^*_h, S^1_h, \ldots, S^{L^*_h}_h\}$. For the root $R^*_h = P_i$ of each star $B^*_h$, we select an arbitrary vertex $r^*_h \in P_i$. Let $(B^*_{h,j})_{h \in [k]}$ be the following collection of stars (which are mutually induced) on $G$: $B^*_{h,j} = \{r_h^* \cup \bigcup_{i=1}^{L^*_h} I(S^i_h)^*\}$. In brief, the collection $(B^*_{h,j})_{h \in [k]}$ is made up of an arbitrary vertex of each root of stars $B^*_h$ and a maximum-weighted independent set of each leaf of $B^*_h$. Remember that we computed this collection of stars for $T_1$; we return a maximum-weighted collection $(B^*_{h,j})_{h \in [k]}$ among all the ones determined for $T_j$, $j \in [2d + 1]$.

**Approximation ratio.** Any collection $B_h$ with stars belonging only to black edges of $G/P$ reveals a collection of stars on the quotient graph. Concretely, two black edges of $G/P$ containing each a branch of $B_h$ must be either non-adjacent or form an induced 3-vertex path on $G' = (V(G/P), E(G/P))$. Conversely, considering a collection $B^*$ of mutually induced
stars of $G'$ and, for each $e \in B^*$, a collection $B^*_e$ of mutually induced stars on $G[p(e)]$ produces a global collection of stars of $G$: then, we can partition its edges into $2d + 1$ parts (as with $T_1, \ldots, T_{2d+1}$) such that each part contains mutually induced stars. As the collection $B$ computed above provides us with a heaviest collection of $G'$, a maximum-weighted $B^*_h$ over all $T_j$ is a $2d + 1$-approximation for $B$, whose weight is at least the weight of $A_b$.

**Conclusion of the proof.** We finally output a heaviest collection of mutually induced stars among the three approximating respectively $A_v$, $A_r$, and $A_b$. The overall running time is in $2^{O_d(\sqrt{n})}$. An upper bound for the approximation ratio of this algorithm is $3z = O(d^2)$. ▶

As for the other problems treated in this article, we apply to **MAX LEAVES INDUCED STAR FOREST** the time-approximation trade-off proposed in Lemma 22.

**Theorem 36.** **MAX LEAVES INDUCED STAR FOREST** on an $n$-vertex graph $G$, weight function $w_V$, with prescribed set $Y \subseteq E(G)$, and given with a $d'$-sequence, satisfies the assumptions of Lemma 22. In particular, this problem admits

$I.$ a $(d + 1)2^{d-1}$-approximation in time $2^{O_d(n \log n)}$, for every integer $q \geq 0$;

$II.$ an $\varepsilon$-approximation in polynomial-time $O_d(1) \log^{O_d(1)} n \cdot n^{O(1)}$, for any $\varepsilon > 0$, and

$III.$ a log $n$-approximation in time $2^{O_d(n^{\frac{1}{\log \log n}})}$,

with $d := c_2d + 2 \cdot 2^{c_3d+\varepsilon} + 4$.

**Proof.** The exhaustive algorithm (trying out all vertex subsets and checking whether they induce a collection of mutually induced stars in $Y$) solves **MAX LEAVES INDUCED STAR FOREST** in time $2^{O(n)}$. Thus we show **MAX LEAVES INDUCED STAR FOREST** satisfies the second item of Lemma 22. We set $c_2 \geq 1$ as the required exponent to turn a $d'$-sequence into a neatly divided matrix of $M_{n, 2d+2}$ conform to $G$, and compute the various needed colorings, the appropriate $\frac{1}{2} < c_1 < 1$, and $2 < c_3 < 3$, and $f(d) = 2d > 1$.

**Approximating $A_v$.** The algorithm makes $\lceil \sqrt{n} \rceil$ recursive calls to solve **MAX LEAVES INDUCED STAR FOREST** on parts $P_i$. Furthermore, $d + 1$ calls to WMIS are needed on induced subgraphs of $G/P$. All of these induced subgraphs are on at most $d\sqrt{n}$ vertices.

**Approximating $A_r$.** The algorithm makes at most $\frac{\sqrt{2d}}{2}$ recursive calls (one call per red edge of $G/P$) on induced subgraphs of $G$ with at most $2d\sqrt{n}$ vertices, followed by at most $2d(d - 1)d + 1$ calls of **MAX SUBSET INDUCED MATCHING** on full cleanups of induced subgraphs of $G/P$ with at most $\sqrt{n}$ vertices.

**Approximating $A_b$.** The algorithm makes $\lceil \sqrt{n} \rceil$ calls to solve WMIS on parts $P_i$ and one recursive call on a full cleanup of $G/P$ on $\lceil \sqrt{n} \rceil$ vertices.

In summary, we make $O_d(\sqrt{n})$ recursive calls or calls to problems WMIS and **MAX SUBSET INDUCED MATCHING** (which already satisfy Lemma 22 with better constants) on induced subgraphs of $G$ or full cleanups of (the whole) $G/P$, each on $O_d(\sqrt{n})$ vertices. Hence, by Lemma 18, the induction applies.

Getting $r$-approximations on every subcall allows us to output a global $3(2(d - 1)d + 1)r^2$-approximation for **MAX LEAVES INDUCED STAR FOREST**:

$I.$ collection $A_v$ is approximated with ratio $(d + 1)r^2$

$II.$ collection $A_r$ is approximated with ratio $(2d - 1)d + 1)r^2$

$III.$ collection $A_b$ is approximated with ratio $(2d + 1)r^2$.

The extra factor 3 comes from the fact that we output the heaviest of these three solutions. ▶

**MAX EDGE INDUCED STAR FOREST** is a particular case of **MAX LEAVES INDUCED STAR FOREST** with $w_V(u) = 1$ for every vertex $u \in V(G)$. Furthermore, a solution of **MAX EDGE INDUCED STAR FOREST** is a 3-approximation of a solution of **MAX EDGE INDUCED FOREST**.

These observations together with Theorem 36 allow us to state the following result.
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Corollary 37. **Max Edge Induced Star Forest** and **Max Edge Induced Forest** on an $n$-vertex graph $G$, with prescribed set $Y \subseteq E(G)$, and given with a $d'$-sequence, admit

- an $n^c$-approximation in polynomial-time $O_{d',c}(1) \log^{O_d(1)} n \cdot n^{O(1)}$, for any $\varepsilon > 0$, and

- a $\log n$-approximation in time $2^{O_d(n^{1-\varepsilon}/\log n)}$,

with $d := c_{2d'+2} \cdot 2^{k_{2d'+2}+4}$.

7 Limits

We now discuss the limits of our framework. We give some examples of problems that are unlikely to have an $n^{1-\varepsilon}$-approximation algorithm on graphs of bounded twin-width. The first such problem is **Min Independent Dominating Set**, where one seeks a minimum-cardinality set which is both an independent set and a dominating set. In general $n$-vertex graphs, this problem cannot be $n^{1-\varepsilon}$-approximated in polynomial time unless $P=NP$ [22], and cannot be $r$-approximated in time $2^{o(n/r)}$ for any $r = r(n)$, unless the ETH fails [11].

We show that **Min Independent Dominating Set** has the same polytime inapproximability in graphs of bounded twin-width.

Theorem 38. **For every** $\varepsilon > 0$, **Min Independent Dominating Set** **cannot be** $n^{1-\varepsilon}$-approximated in polynomial time on $n$-vertex graphs of twin-width at most 9 given with a 9-sequence, unless $P=NP$.

Proof. We perform the classic reduction of Halldórsson from SAT [22], but from **Planar 3-SAT** where each literal has at most two occurrences, which remains NP-complete [29]. More precisely we add a triangle $d_i, t_i, f_i$ for each variable $x_i$ (with $i \in [N]$), and an independent set $I_j$ of size $r$ for each 3-clause $C_j$ (with $j \in [M]$). We link $t_i$ to all the vertices of $I_j$ whenever $x_i$ appears positively in $C_j$, and we link $f_i$ to all the vertices of $I_j$ whenever $x_i$ appears negatively in $C_j$. This defines a graph $G$ with $n = 3N + rM$ vertices.

It can be observed that if the Planar 3-SAT instance is satisfiable, then there is an independent dominating set of size $N$, whereas if the formula is unsatisfiable then any independent dominating set has size at least $r$. Setting $r := N^{1/2}$, the gap between positive and negative instances is $\Theta(1) n^{1-\epsilon}$, while preserving the fact that the reduction is polynomial.

Let us now argue that $G$ has twin-width at most 9, and that a 9-sequence of it can be computed in polynomial time. We can first contract each $I_j$ into a single vertex without creating a red edge. Next we can contract every triangle $d_i, t_i, f_i$ into a single vertex of red degree at most 4. At this point, the current trigraph is a planar graph of maximum degree at most 4. It was observed in [9] that planar trigraphs with maximum (total) degree at most 9 have twin-width at most 9. This is because any planar graph has a pair of vertices on the same face with at most 9 neighbors (outside of themselves) combined [28]. Hence we get a 9-sequence for $G$ that can be computed in polynomial time. Incidentally the twin-width of planar graphs (that is, planar trigraphs without red edge) but no restriction on the maximum degree is also at most 9 [24].

Another very inapproximable is **Longest Induced Path**, which also does not admit a polytime $n^{1-\varepsilon}$-approximation algorithm unless $P=NP$ [31], and cannot be $r$-approximated in time $2^{o(n/r)}$ for any $r = o(n)$, unless the ETH fails [11]. The non-induced version, the **Longest Path** problem, has a notoriously big gap between the best known approximation algorithm whose factor is $n/\exp(\Omega(\sqrt{\log n}))$ [18], and the sharpest conditional lower bound which states that, for any $\varepsilon > 0$, a $2^{\log^{1-\varepsilon} n}$-approximation would imply that $NP \subseteq QP$ [27].
Despite being an open question for decades the existence or conditional impossibility of an approximation algorithm for Longest Path with approximation factor, say, $\sqrt{n}$ has not been settled. Nor do we know whether an $n^\varepsilon$-approximation for any $\varepsilon > 0$ is possible. We now show that using our framework to obtain an $n^\varepsilon$-approximation for Longest Induced Path of Longest Path in graphs of bounded twin-width is unlikely to work, in the sense that it would immediately yield such an approximation factor for Longest Path in general graphs.

**Theorem 39.** For any $r = \omega(1)$, an $r$-approximation for Longest Induced Path or Longest Path on graphs of twin-width at most 4 given with a 4-sequence would imply a $(1 + o(1))r$-approximation for Longest Path in general graphs.

**Proof.** It was shown in [3] that any graph obtained by subdividing every edge of an $n$-vertex graph at least $2\log n$ has twin-width at most 4. Besides, a 4-sequence can then be computed in polynomial time.

Let $G$ be any graph with minimum degree at least 2 (note that this restriction does not make Longest Path easier to approximate), and $G'$ be obtained from $G$ by subdividing each of its edges $2\lceil \log n \rceil$ times, and let $s := 2\lceil \log n \rceil + 1$. Let us observe that $G$ has a path of length $\ell$ if and only if $G'$ has a path of length $(\ell + 2)s - 2$ if and only if $G'$ has an induced path of length $(\ell + 2)s - 4$. Hence a polytime $r$-approximation for Longest Induced Path or Longest Path in graphs of bounded twin-width given a 4-sequence would translate into a $(1 + o(1))r$-approximation for Longest Path in general graphs. 

We can use Theorem 39 to get a similar weak obstruction to an $n^\varepsilon$-approximation for Mutually Induced $\mathcal{H}$-packing in graphs of bounded twin-width, for some infinite family of connected graphs $\mathcal{H}$. Recall that by Lemma 34 such an approximation algorithm does exist when $\mathcal{H}$ is a finite collection of connected graphs.

Setting $\mathcal{H}$ to be the set of all paths does not serve that purpose, since one can then use the approximation algorithm for Max Induced Matching. Nevertheless this almost works. We just need to decorate the endpoints of the paths. For every positive integer $n$, let $D_n$ be the decorated path of length $n$, obtained from the $n$-vertex path $P_n$ by adding for each endpoint $u$ two adjacent vertices $u', u''$ both adjacent to $u$. Informally, $D_n$ is a path terminated by a triangle at each end.

**Theorem 40.** Let $\mathcal{H} := \{D_n : n \in \mathbb{N}^+\}$ be the family of all decorated paths. If for every $\varepsilon > 0$, Mutually Induced $\mathcal{H}$-packing admits an $n^\varepsilon$ on $n$-vertex graphs of bounded twin-width given with a 4-sequence, then so does Longest Path on general graphs.

**Proof.** Let $G$ be any graph. For every pair $u \neq v \in V(G)$, define $G_{uv}$ as the graph obtained from $G$ by subdividing all its edges $2\lceil \log(n+2) \rceil$ times, and adding two adjacent vertices $u', u''$ both adjacent to $u$, and two adjacent vertices $v', v''$ both adjacent to $v$. Since there are only two triangles in $G_{uv}$, only one graph of $\mathcal{H}$ can be present in a (mutually induced) packing. Thus Mutually Induced $\mathcal{H}$-packing is now equivalent to finding a longest path between $u$ and $v$. An $n^\varepsilon$-approximation algorithm for this problem would, by Theorem 39, give a similar approximation algorithm for Longest Path in general graphs.

Despite $u', u'', v', v''$, $G_{uv}$ still admits a 4-sequence. For instance, first contract $u'$ and $u''$, and contract $v'$ and $v''$; this does not create red edges, and has the same effect as deleting $u''$ and $v''$. The obtained graph is an induced subgraph of a $2\lceil \log(n+2) \rceil$-subdivision (of a graph on at most $n+2$ vertices). Hence it admits a polytime computable 4-sequence [3].
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